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LINUX INSIDE

на русском

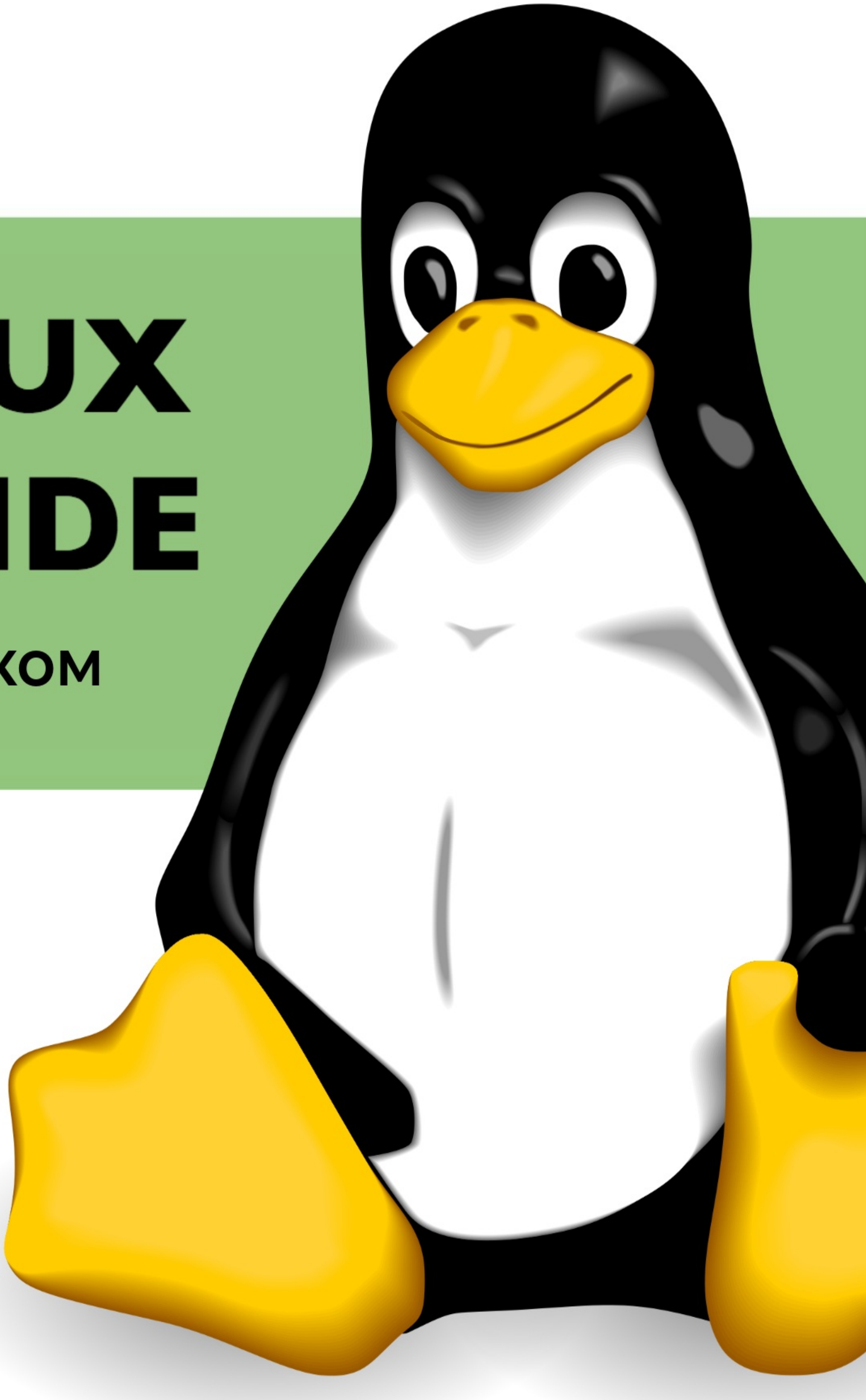

Table of Contents

Содержание

Linux Inside на русском

Оригинальный проект

Серия статей о ядре Linux и его внутреннем устройстве.

Цель проста (от автора) - поделиться своим скромным знанием о внутренностях ядра Linux и помочь людям, которые заинтересованы низкоуровневыми подробостями Linux.

Вопросы/Предложения по переводу: Не стесняйтесь задавать любые вопросы и предложения мне на email.

Электронные версии книги: вы можете найти PDF/EPUB/MOBI в GitBook или в папке репозитория Ebooks (периодически обновляется). Также вы можете сгенерировать их самостоятельно.

Поддержка

Если вам нравится `linux-insides-ru`, вы можете поддержать **автора перевода** следующими способами. Если у вас возникли проблемы с оплатой или вы хотите пожертвовать иным способом, свяжитесь со мной: `proninyaroslav@mail.ru`. Спасибо!

- **Yandex Money**: `410011738561939`
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Такkже вы можете поддержать **оригинального автора** следующими способами:

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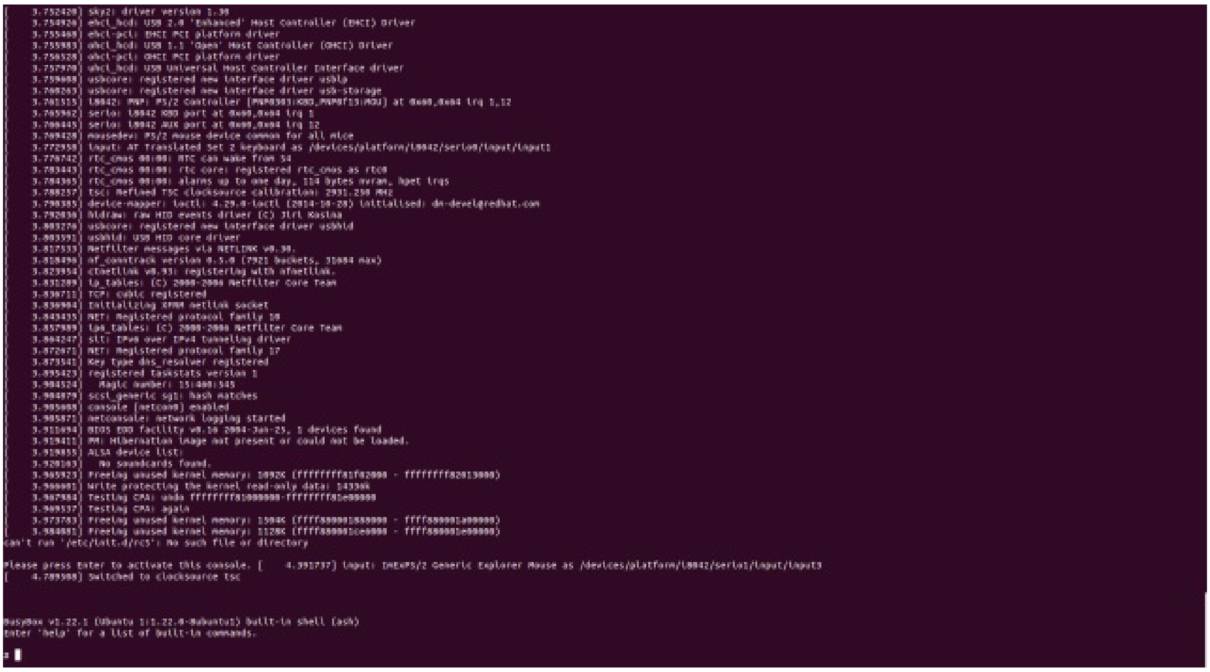
Если у Вас возникнут какие-либо проблемы, не стесняйтесь создавать issues или pull-request.

Пожалуйста, прочтите CONTRIBUTING.md, прежде чем внести изменения.



Список рассылки

Оригинальный проект имеет список рассылки Google Group для изучения исходного кода ядра. Вот несколько инструкций о том, как его использовать.

Присоединиться к списку

Отправьте письмо с любым вопросом/содержимым по адресу `kernelhacking+subscribe@googlegroups.com` . Затем вы получите подтверждение по электронной почте. Ответьте на него с любым содержимым.

> Если у вас есть учётная запись Google, вы также можете открыть страницу архива и нажать **Чтобы отправлять сообщения, вступите в группу**. Вы будете автоматически добавлены в список рассылки.

Отправка писем в список рассылки

Просто отправляйте письма по адресу `kernelhacking@googlegroups.com` . Основы использование те же самые, что и у других списков рассылок, работающих на mailman.

Архив

https://groups.google.com/forum/#!forum/kernelhacking

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Лицензия

Процесс загрузки ядра

Эта глава описывает процесс загрузки ядра Linux. Здесь вы увидите серию статей, которые описывают полный цикл загрузки ядра:

- От загрузчика к ядру - описывает все стадии от включения компьютера до запуска первой инструкции ядра.
- Первые шаги в коде настройки ядра - описывает первые шаги в коде настройки ядра. Вы увидите инициализацию динамической памяти, запросы различных параметров, таких как EDD, IST и др.
- Инициализация видеорежима и переход в защищённый режим - описывает инициализацию видеорежима в коде настройки ядра и переход в защищённый режим.
- Переход в 64-битный режим - описывает подготовку к переходу в 64-битный режим и детали перехода.
- Декомпрессия ядра - описывает подготовку перед декомпрессией ядра и детали самой декомпрессии.
- Рандомизация адреса ядра - описывает рандомизацию адреса загрузки ядра Linux.

Эта глава соответствует `ядру Linux v4.17` .

Процесс загрузки ядра. Часть 1.

От загрузчика к ядру

От автора:

Если вы читали предыдущие статьи моего блога, то могли заметить, что некоторое время назад я начал увлекаться низкоуровневым программированием. Я написал несколько статей о программировании под x86_64 в Linux. В то же время я начал "погружаться" в исходный код Linux. Я испытываю большой интерес к пониманию того, как работают низкоуровневые вещи, как запускаются программы на моем компьютере, как они расположены в памяти, как ядро управляет процессами и памятью, как работает сетевой стек на низком уровне и многие другие вещи. Поэтому я решил написать ещё одну серию статей о ядре Linux для **x86_64**.

Замечу, что я не профессиональный хакер ядра и не пишу под него код на работе. Это просто хобби. Мне просто нравятся низкоуровневые вещи и мне интересно наблюдать за тем, как они работают. Так что, если вас что-то будет смущать или у вас появятся вопросы или замечания, пишите мне в твиттер 0xAX, присылайте письма на email или просто создавайте issue. Я ценю это.

Все статьи также будут доступны в репозитории Github, и, если вы обнаружите какую-нибудь ошибку в содержимом статьи, не стесняйтесь присылать pull request.

Заметьте, что это не официальная документация, а просто материал для обучения и обмена знаниями.

Требуемые знания

- Понимание кода на языке C
- Понимание кода на языке ассемблера (AT&T синтаксис)

В любом случае, если вы только начинаете изучать подобные инструменты, я постараюсь объяснить некоторые детали в этой и последующих частях. Ладно, простое введение закончилось, и теперь можно начать "погружение" в ядро и низкоуровневые вещи.

Я начал писать эту книгу, когда актуальной версией ядра Linux была 3.18, и с этого момента многое могло измениться. При возникновении изменений я буду соответствующим образом обновлять статьи.

Магическая кнопка включения, что происходит дальше?

Несмотря на то, что это серия статей о ядре Linux, мы не будем начинать непосредственно с его исходного кода (по крайней мере в этом параграфе). Как только вы нажмёте магическую кнопку включения на вашем ноутбуке или настольном компьютере, он начинает работать. Материнская плата посылает сигнал к источнику питания. После получения сигнала, источник питания обеспечивает компьютер надлежащим количеством электричества. После того как материнская плата получает сигнал "питание в норме" (Power good signal), она пытается запустить CPU. CPU сбрасывает все остаточные данные в регистрах и записывает предустановленные значения каждого из них.

CPU серии Intel 80386 и старше после перезапуска компьютера заполняют регистры следующими предустановленными значениями:

```
IP          0xfff0
CS selector 0xf000
CS base     0xffff0000
```

Процессор начинает свою работу в режиме реальных адресов. Давайте немного задержимся и попытаемся понять сегментацию памяти в этом режиме. Режим реальных адресов поддерживается всеми x86-совместимыми процессорами, от 8086 до самых новых 64-битных CPU Intel. Процессор 8086 имел 20-битную шину адреса, т.е. он мог работать с адресным пространством в диапазоне 0-0xFFFFF (1 мегабайт). Но регистры у него были только 16-битные, а в таком случае максимальный размер адресуемой памяти составляет `2^16 - 1` или `0xffff` (64 килобайта).

Сегментация памяти используется для того, чтобы задействовать всё доступное адресное пространство. Вся память делится на небольшие, фиксированного размера сегменты по `65536` байт (64 Кб) каждый. Поскольку мы не можем адресовать память свыше `64 Кб` с помощью 16-битных регистров, был придуман альтернативный метод.

Адрес состоит из двух частей: селектора сегмента, который содержит базовый адрес, и смещение от этого базового адреса. В режиме реальных адресов базовый адрес селектора сегмента это `Селектор Сегмента * 16`. Таким образом, чтобы получить физический адрес в памяти, нужно умножить селектор сегмента на `16` и прибавить к нему смещение:

```
Физический адрес = Селектор сегмента * 16 + Смещение
```

Например, если `CS:IP` содержит `0x2000:0x0010`, то соответствующий физический адрес будет:

```
>>> hex((0x2000 << 4) + 0x0010)
'0x20010'
```

Но если мы возьмём максимально доступный селектор сегментов и смещение: `0xffff:0xffff`, то итоговый адрес будет:

```
>>> hex((0xffff << 4) + 0xffff)
'0x10ffef'
```

что больше первого мегабайта на 65520 байт. Поскольку в режиме реальных адресов доступен только один мегабайт, с отключённой адресной линией A20 `0x10ffef` становится `0x00ffef`.

Хорошо, теперь мы немного знаем о режиме реальных адресов и адресации памяти в этом режиме. Давайте вернёмся к обсуждению значений регистров после сброса.

Регистр `CS` состоит из двух частей: видимый селектор сегмента и скрытый базовый адрес. В то время как базовый адрес, как правило, формируется путём умножения значения селектора сегмента на 16, во время аппаратного перезапуска в селектор сегмента в регистре `CS` записывается `0xf000`, а в базовый адрес - `0xffff0000`. Процессор использует этот специальный базовый адрес, пока регистр `CS` не изменится.

Начальный адрес формируется путём добавления базового адреса к значению в регистре `EIP`:

```
>>> 0xffff0000 + 0xfff0
'0xfffffff0'
```

Мы получили `0xfffffff0`, т.е. 16 байт ниже 4 Гб. По этому адресу располагается так называемый вектор прерываний. Это область памяти, в которой CPU ожидает найти первую инструкцию для выполнения после сброса. Она содержит инструкцию jump (`jmp`), которая обычно указывает на точку входа в BIOS. Например, если мы взглянем на исходный код coreboot (`src/cpu/x86/16bit/reset16.inc`), то увидим следующее:

```
    .section ".reset", "ax", %progbits
    .code16
.globl  _start
_start:
    .byte  0xe9
    .int   _start16bit - ( . + 2 )
    ...
```

Здесь мы можем видеть опкод инструкции jmp - `0xe9`, и его адрес назначения `_start16bit - ( . + 2)`.

Мы также можем видеть, что секция `reset` занимает `16` байт и начинается с `0xfffffff0` (`src/cpu/x86/16bit/reset16.ld`):

```
SECTIONS {
    /* Trigger an error if I have an unuseable start address */
    _bogus = ASSERT(_start16bit >= 0xffff0000, "_start16bit too low. Please report.");
    _ROMTOP = 0xfffffff0;
    . = _ROMTOP;
    .reset . : {
        *(.reset);
        . = 15;
        BYTE(0x00);
    }
}
```

Теперь запускается BIOS; после инициализации и проверки оборудования, BIOS нужно найти загрузочное устройство. Порядок загрузки хранится в конфигурации BIOS, которая определяет, с каких устройств BIOS пытается загрузиться. При попытке загрузиться с жёсткого диска, BIOS пытается найти загрузочный сектор. На размеченных жёстких дисках со схемой разделов MBR, загрузочный сектор расположен в первых 446 байтах первого сектора, размер которого 512 байт. Последние два байта первого сектора - `0x55` и `0xaa`, которые оповещают BIOS о том, что устройство является загрузочным.

Например:

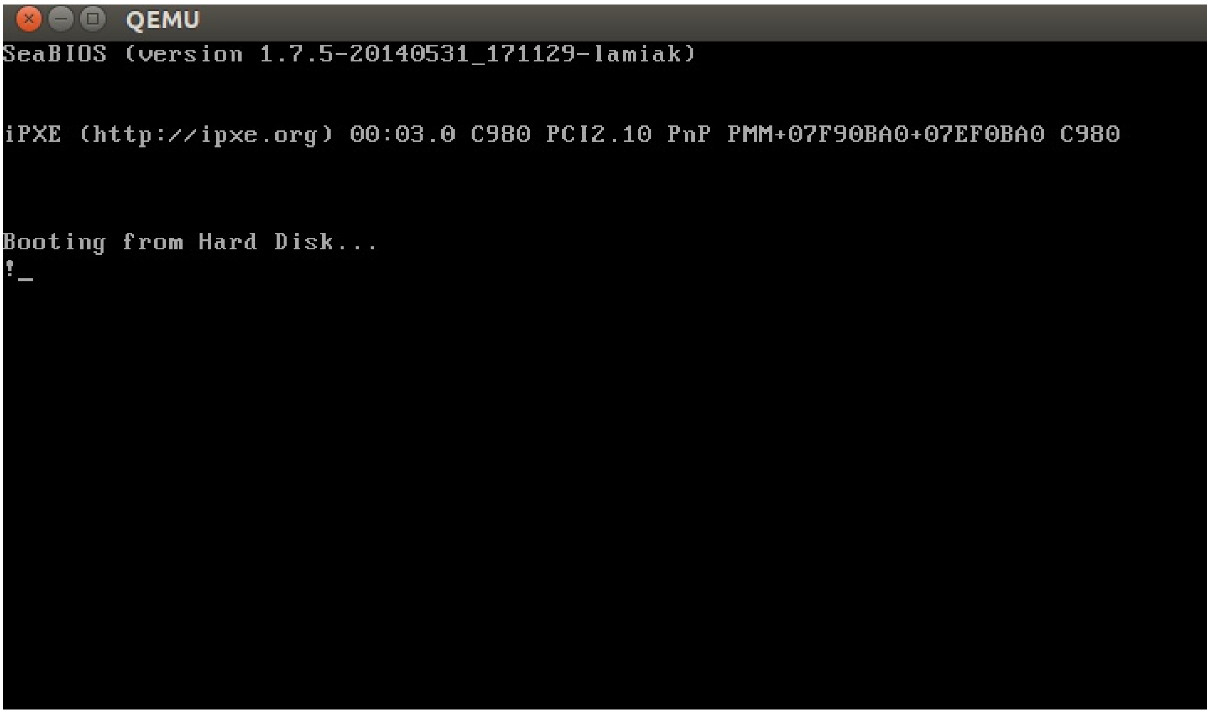
```
;
; Замечание: этот пример написан с использованием Intel синтаксиса
;
[BITS 16]

boot:
    mov al, '!'
    mov ah, 0x0e
    mov bh, 0x00
    mov bl, 0x07

    int 0x10
    jmp $

times 510-($-$$) db 0

db 0x55
db 0xaa
```

Собрать и запустить этот код можно таким образом:

```
nasm -f bin boot.nasm && qemu-system-x86_64 boot
```

Команда оповещает эмулятор QEMU о необходимости использовать в качестве образа диска созданный только что бинарный файл. Пока последний проверяет, удовлетворяет ли загрузочный сектор всем необходимым требованиям (в origin записывается `0x7c00`, а в конце магическая последовательность), QEMU будет работать с бинарным файлом как с главной загрузочной записью (MBR) образа диска.

Вы увидите:



В этом примере мы можем видеть, что код будет выполнен в 16-битном режиме реальных адресов и начнёт выполнение с адреса `0x7c00`. После запуска он вызывает прерывание 0x10, которое просто печатает символ `!`. Оставшиеся 510 байт заполняются нулями, и код заканчивается двумя магическими байтами `0xaa` и `0x55`.

Вы можете увидеть бинарный дамп с помощью утилиты `objdump`:

```
nasm -f bin boot.nasm
objdump -D -b binary -mi386 -Maddr16,data16,intel boot
```

Реальный загрузочный сектор имеет код для продолжения процесса загрузки и таблицу разделов вместо кучи нулей и восклицательного знака :) С этого момента BIOS передаёт управление загрузчику.

ЗАМЕЧАНИЕ: Как уже было упомянуто ранее, CPU находится в режиме реальных адресов; в этом режиме вычисление физического адреса в памяти выполняется следующим образом:

```
Физический адрес = Селектор сегмента * 16 + Смещение
```

так же, как было упомянуто выше. У нас есть только 16-битные регистры общего назначения; максимальное значение 16-битного регистра - `0xffff`. Поэтому, если мы возьмём максимальное возможное значение, то результат будет следующий:

```
>>> hex((0xffff * 16) + 0xffff)
'0x10ffef'
```

где `0x10ffef` равен `1 Мб + 64 Кб - 16 байт`. Процессор 8086 (который был первым процессором с режимом реальных адресов), в отличии от этого, имеет 20-битную шину адресации. Поскольку `2^20 = 1048576` это 1 Мб, получается, что фактический объём доступной памяти составляет 1 Мб.

В целом, карта разделов в режиме реальных адресов выглядит следующим образом:

```
0x00000000 - 0x000003FF - Таблица векторов прерываний
0x00000400 - 0x000004FF - Данные BIOS
0x00000500 - 0x00007BFF - Не используется
0x00007C00 - 0x00007DFF - Наш загрузчик
```

```
0x00007E00 - 0x0009FFFF - Не используется
0x000A0000 - 0x000BFFFF - RAM (VRAM) видеопамять
0x000B0000 - 0x000B7777 - Память для монохромного видео
0x000B8000 - 0x000BFFFF - Память для цветного видео
0x000C0000 - 0x000C7FFF - BIOS ROM видеопамяти
0x000C8000 - 0x000EFFFF - Скрытая область BIOS
0x000F0000 - 0x000FFFFF - Системная BIOS
```

В начале статьи я написал, что первая инструкция, выполняемая CPU, расположена по адресу `0xFFFFFFF0`, значение которого намного больше, чем `0xFFFFF` (1 Мб). Каким образом CPU получает доступ к этому участку в режиме реальных адресов? Ответ на этот вопрос находится в документации coreboot:

```
0xFFFE_0000 - 0xFFFF_FFFF: 128 Кб ROM отображаются на адресное пространство
```

В начале выполнения BIOS находится не в RAM, а в ROM.

Загрузчик

Существует несколько загрузчиков Linux, такие как GRUB 2 и syslinux. Ядро Linux имеет протокол загрузки, который определяет требования к загрузчику для реализации поддержки Linux. В этом примере будет описан GRUB 2.

Теперь, когда BIOS выбрал загрузочное устройство и передал контроль управления коду в загрузочном секторе, начинается выполнение boot.img. Этот код очень простой в связи с ограниченным количеством свободного пространства и содержит указатель, который используется для перехода к основному образу GRUB 2. Основной образ начинается с diskboot.img, который обычно располагается сразу после первого сектора в неиспользуемой области перед первым разделом. Приведённый выше код загружает оставшуюся часть основного образа, который содержит ядро и драйверы GRUB 2 для управления файловой системой. После загрузки остальной части основного образа, код выполняет функция grub_main.

Функция `grub_main` инициализирует консоль, получает базовый адрес для модулей, устанавливает корневое устройство, загружает/обрабатывает файл настроек grub, загружает модули и т.д. В конце выполнения, `grub_main` переводит grub обратно в нормальный режим. Функция `grub_normal_execute` (из `grub-core/normal/main.c`) завершает последние приготовления и отображает меню выбора операционной системы. Когда мы выбираем один из пунктов меню, запускается функция `grub_menu_execute_entry`, которая в свою очередь запускает команду grub `boot`, загружающую выбранную операционную систему.

Из протокола загрузки видно, что загрузчик должен читать и заполнять некоторые поля в заголовке ядра, который начинается со смещения `0x01f1` в коде настроек. Вы можете посмотреть загрузочный скрипт компоновщика, чтобы убедиться в этом. Заголовок ядра arch/x86/boot/header.S начинается с:

```
    .globl hdr
hdr:
    setup_sects: .byte 0
    root_flags:  .word ROOT_RDONLY
    syssize:     .long 0
    ram_size:    .word 0
    vid_mode:    .word SVGA_MODE
    root_dev:    .word 0
    boot_flag:   .word 0xAA55
```

Загрузчик должен заполнить этот и другие заголовки (которые помеченные как тип `write` в протоколе загрузки Linux, например, в данном примере) значениями, которые он получил из командной строки или значениями, вычисленными во время загрузки. (Мы не будет вдаваться в подробности и описывать все поля заголовка ядра, но мы сделаем это, когда будем обсуждать как их использует ядро; тем не менее вы можете найти полное описание всех полей в протоколе загрузки.)

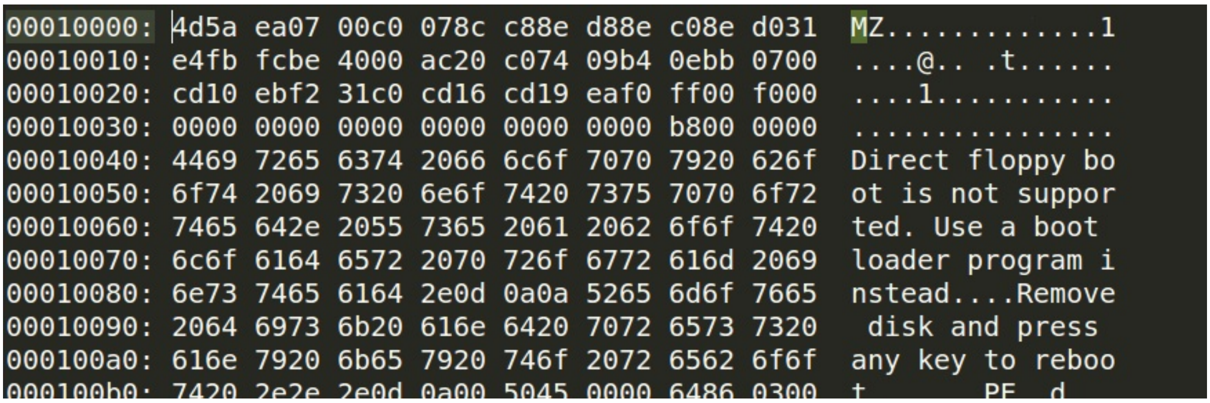
Как мы видим из протокола, после загрузки ядра карта распределения памяти будет выглядеть следующим образом:

```
         | Ядро защищённого режима  |
100000   +--------------------------+
         | Память I/O               |
0A0000   +--------------------------+
         | Резерв для BIOS          | Оставлен максимально допустимый размер
         ~                          ~
         | Командная строка         | (Может быть ниже X+10000)
X+10000  +--------------------------+
         | Стек/куча                | Используется кодом ядра в режиме реальных адресов
X+08000  +--------------------------+
         | Настройки ядра           | Код ядра режима реальных адресов.
         | Загрузочный сектор ядра  | Унаследованный загрузочный сектор ядра.
       X +--------------------------+
         | Загрузчик                | <- Точка входа загрузочного сектора 0x7C00
001000   +--------------------------+
         | Резерв для MBR/BIOS      |
000800   +--------------------------+
         | Обычно используется MBR  |
000600   +--------------------------+
         | Используется только BIOS |
000000   +--------------------------+
```

Итак, когда загрузчик передаёт управление ядру, он запускается с:

```
X + sizeof(KernelBootSector) + 1
```

где `X` - это адрес загруженного сектора загрузки ядра. В моем случае `X` это `0x10000`, как мы можем видеть в дампе памяти:


```
00010000: 4d5a ea07 00c0 078c c88e d88e c08e d031  MZ.............1
00010010: e4fb fcbe 4000 ac20 c074 09b4 0ebb 0700  ....@.. .t......
00010020: cd10 ebf2 31c0 cd16 cd19 eaf0 ff00 f000  ....1...........
00010030: 0000 0000 0000 0000 0000 0000 b800 0000  ................
00010040: 4469 7265 6374 2066 6c6f 7070 7920 626f  Direct floppy bo
00010050: 6f74 2069 7320 6e6f 7420 7375 7070 6f72  ot is not suppor
00010060: 7465 642e 2055 7365 2061 2062 6f6f 7420  ted. Use a boot
00010070: 6c6f 6164 6572 2070 726f 6772 616d 2069  loader program i
00010080: 6e73 7465 6164 2e0d 0a0a 5265 6d6f 7665  nstead....Remove
00010090: 2064 6973 6b20 616e 6420 7072 6573 7320   disk and press
000100a0: 616e 7920 6b65 7920 746f 2072 6562 6f6f  any key to reboo
000100b0: 7420 2e2e 2e0d 0a00 5045 0000 6486 0300  t.......PE..d...
```

Сейчас загрузчик поместил ядро Linux в память, заполнил поля заголовка, а затем переключился на него. Теперь мы можем перейти непосредственно к коду настройки ядра.

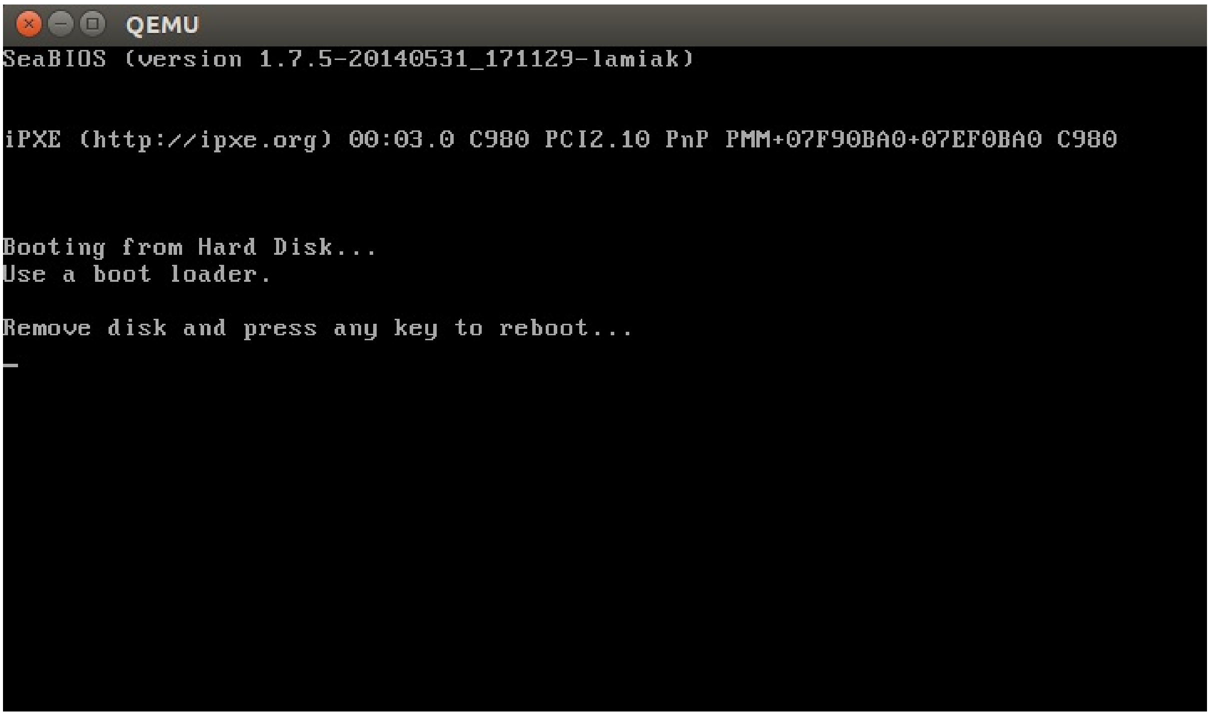
Начальный этап настройки ядра

Наконец, мы находимся в ядре! Технически, ядро ещё не работает; во-первых, часть ядра, отвественная за настройку, должна подготовить декомпрессор, вещи связанные с управлением памятью и т.д. После всех подготовок код настройки ядра должен распаковывать фактическое ядро и совершить переход на него. Выполнение настройки начинается в файле arch/x86/boot/header.S, начиная с метки _start. Это немного странно на первый взгляд, так как перед этим есть ещё несколько инструкций.

Давным-давно у Linux был свой загрузчик, но сейчас, если вы запустите, например:

```
qemu-system-x86_64 vmlinuz-3.18-generic
```

то увидите:



На самом деле, `header.S` начинается с MZ (см. картинку выше), вывода сообщения об ошибке и PE заголовка:

```
#ifdef CONFIG_EFI_STUB
# "MZ", MS-DOS header
.byte 0x4d
.byte 0x5a
#endif
...
...
...
pe_header:
    .ascii "PE"
    .word 0
```

Это нужно, чтобы загрузить операционную систему с поддержкой UEFI. Мы не будем рассматривать его внутреннюю работу прямо сейчас; мы сделаем это в одной из следующих глав.

Настоящая настройка ядра начинается с:

```
// header.S строка 292
.globl _start
_start:
```

Загрузчик (grub2 или другой) знает об этой метке (при смещении `0x200` от `MZ`) и сразу переходит на неё, несмотря на то, что `header.S` начинается с секции `.bstext` , который выводит сообщение об ошибке:

```
//
// arch/x86/boot/setup.ld
//
. = 0;                    // текущая позиция
.bstext : { *(.bstext) }  // поместить секцию .bstext в позицию 0
.bsdata : { *(.bsdata) }
```

Точка входа настройки ядра:

```
    .globl _start
_start:
    .byte  0xeb
    .byte  start_of_setup-1f
1:
    //
    // остальная часть заголовка
    //
```

Здесь мы можем видеть опкод инструкции `jmp` (`0xeb`) к метке `start_of_setup-1f` . Нотация `Nf` означает, что `2f` ссылается на следующую локальную метку `2:` ; в нашем случае это метка `1` , которая расположена сразу после инструкции jump и содержит оставшуюся часть заголовка. Сразу после заголовка настроек мы видим секцию `.entrytext` , которая начинается с метки `start_of_setup` .

Это первый код, который на самом деле запускается (отдельно от предыдущей инструкции jump, конечно). После того как настройщик ядра получил управление от загрузчика, первая инструкция `jmp` располагается по смещению `0x200` от начала реальных адресов, т.е после первых 512 байт. Об этом можно узнать из протокола загрузки ядра Linux, а также увидеть в исходном коде grub2:

```
segment = grub_linux_real_target >> 4;
state.gs = state.fs = state.es = state.ds = state.ss = segment;
state.cs = segment + 0x20;
```

В моём случае, ядро загружается по адресу `0x10000` . Это означает, что после начала настройки ядра регистры сегмента будут иметь следующие значения:

```
gs = fs = es = ds = ss = 0x1000
cs = 0x1020
```

После перехода на метку `start_of_setup` , необходимо соблюсти следующие условия:

- Убедиться, что все значения всех сегментных регистров равны
- Правильно настроить стек, если это необходимо
- Настроить BSS
- Перейти к C-коду в main.c

Давайте посмотрим, как эти условия выполняются.

Выравнивание сегментных регистров

Прежде всего, ядро гарантирует, что сегментные регистры `ds` и `es` указывают на один и тот же адрес. Затем оно сбрасывает флаг направления с помощью инструкции `cld` :

```
    movw    %ds, %ax
    movw    %ax, %es
    cld
```

Как я уже писал ранее, `grub2` загружает код настройки ядра по адресу `0x1000` (адрес по умолчанию) и `cs` по адресу `0x1020` , потому что выполнение не начинается с начала файла, а с метки `_start` :

```
_start:
    .byte 0xeb
    .byte start_of_setup-1f
```

расположеной в `512` байтах от 4d 5a. Нам также необходимо выровнять `cs` с `0x1020` на `0x1000` и остальные сегментные регистры. После этого мы настраиваем стек:

```
    pushw   %ds
    pushw   $6f
    lretw
```

кладём значение `ds` в стек по адресу метки 6 и выполняем инструкцию `lretw`. Когда мы вызываем `lretw`, она загружает адрес метки `6` в регистр счётчика команд (IP), и загружает `cs` со значением `ds`. После этого `ds` и `cs` будут иметь одинаковые значения.

Настройка стека

Почти весь код настройки - это подготовка для среды языка C в режиме реальных адресов. Следующим шагом является проверка значения регистра `ss` и создание корректного стека, если значение `ss` неверно:

```
    movw    %ss, %dx
    cmpw    %ax, %dx
    movw    %sp, %dx
    je      2f
```

Это может привести к трём различны сценариям:

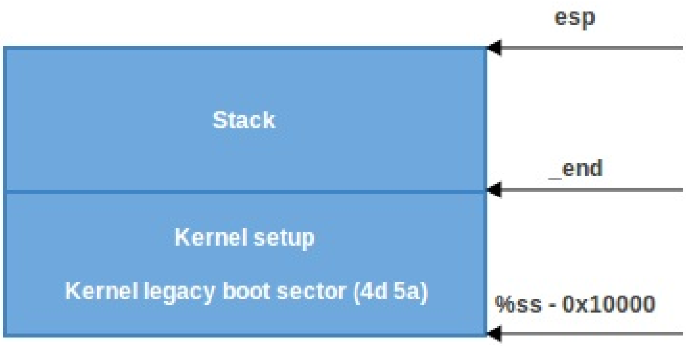
- `ss` имеет верное значение `0x1000` (как и все остальные сегментные регистры рядом с `cs`)
- `ss` является некорректным и установлен флаг `CAN_USE_HEAP` (см. ниже)
- `ss` является некорректным и флаг `CAN_USE_HEAP` не установлен (см. ниже)

Давайте рассмотрим все три сценария:

- `ss` имеет верный адрес (`0x1000`). В этом случае мы переходим на метку 2:

```
2:  andw    $~3, %dx
    jnz     3f
    movw    $0xfffc, %dx
3:  movw    %ax, %ss
    movzwl  %dx, %esp
    sti
```

Здесь мы видим выравнивание сегмента `dx` (содержащего значение `sp`, полученное загрузчиком) до 4 байт и проверку - является ли полученное значение нулём. Если ноль, то помещаем `0xfffx` (выровненный до `4` байт адрес до максимального значения сегмента в 64 Кб) в `dx`. Если не ноль, продолжаем использовать `sp`, полученный от загрузчика (в моём случае `0xf7f4`). После этого мы помещаем значение `ax` в `ss`, который хранит корректный адрес сегмента `0x1000` и устанавливает корректное значение `sp`. Теперь мы имеем корректный стек:



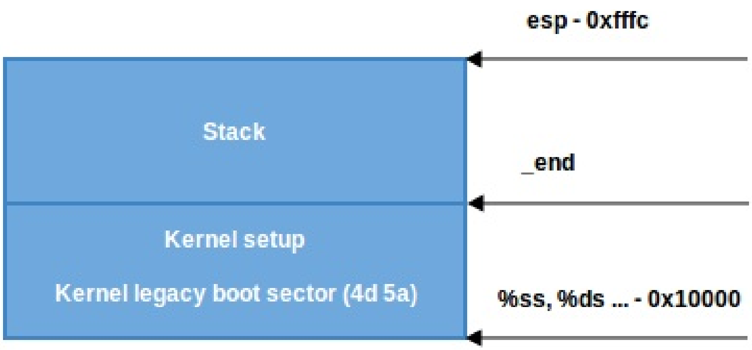
- Второй сценарий (когда `ss` != `ds`). Во-первых, помещаем значение _end (адрес окончания кода настройки) в `dx` и проверяем поле заголовка `loadflags` инструкцией `testb` , чтобы понять, можем ли мы использовать кучу (heap). loadflags является заголовком с битовой маской, который определён как:

```
#define LOADED_HIGH     (1<<0)
#define QUIET_FLAG      (1<<5)
#define KEEP_SEGMENTS   (1<<6)
#define CAN_USE_HEAP    (1<<7)
```

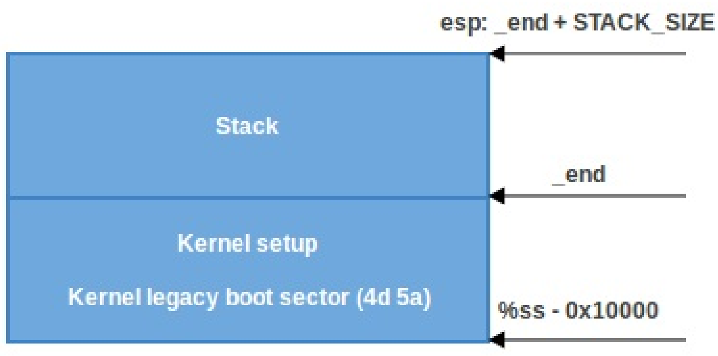
и, как мы можем узнать из протокола загрузки:

```
Имя поля: loadflags

  Данное поле является битовой маской.

  Бит 7 (запись): CAN_USE_HEAP
    Бит, установленный в 1, указывает на корректность heap_end_ptr.
    Если поле очищено, то некоторый функционал кода настройки будет отключен.
```

Если бит `CAN_USE_HEAP` установлен, мы помещаем `heap_end_ptr` в `dx` (который указывает на `_end`) и добавляем к нему `STACK_SIZE` (минимальный размер стека, `512` байт). После этого, если `dx` без переноса (будет без переноса, поскольку `dx = _end + 512`), переходим на метку `2` (как в предыдущем случае) и создаём корректный стек.



- Если флаг `CAN_USE_HEAP` не установлен, мы просто используем минимальный стек от `_end` до `_end + STACK_SIZE` :



Настройка BSS

Последние два шага, которые нужно выполнить перед тем, как мы сможем перейти к основному коду на C, это настройка BSS и проверка "магических" чисел. Сначала проверка чисел:

```
cmpl    $0x5a5aaa55, setup_sig
jne     setup_bad
```

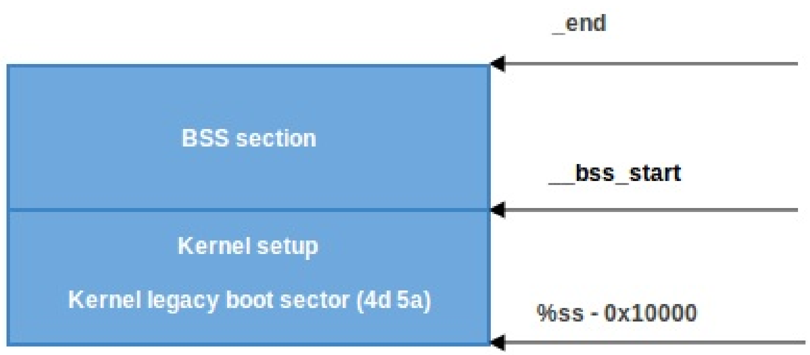
Это простое сравнение setup_sig с магическим числом `0x5a5aaa55`. Если они не равны, сообщается о фатальной ошибке.

Если магические числа совпадают, зная, что у нас есть набор правильно настроенных сегментных регистров и стек, нам всего лишь нужно настроить BSS.

Секция BSS используется для хранения статически выделенных, неинициализированных данных. Linux тщательно обнуляет эту область памяти, используя следующий код:

```
movw    $__bss_start, %di
movw    $_end+3, %cx
xorl    %eax, %eax
subw    %di, %cx
shrw    $2, %cx
rep; stosl
```

Во-первых, адрес __bss_start помещается в `di`. Далее, адрес `_end + 3` (+3 - выравнивает до 4 байт) помещается в `cx`. Регистр `eax` очищается (с помощью инструкции `xor`), а размер секции BSS (`cx` - `di`) вычисляется и помещается в `cx`. Затем `cx` делится на 4 (размер 'слова' (англ. word)), а инструкция `stosl` используется повторно, сохраняя значение `eax` (ноль) в адрес, на который указывает `di`, автоматически увеличивая `di` на 4 (это продолжается до тех пор, пока `cx` не достигнет нуля). Эффект от этого кода в том, что теперь все 'слова' в памяти от `__bss_start` до `_end` заполнены нулями:



Переход к основному коду

Вот и все, теперь у нас есть стек и BSS, поэтому мы можем перейти к C-функции `main()` :

```
    calll main
```

Функция `main()` находится в файле arch/x86/boot/main.c. О том, что она делает, вы сможете узнать в следующей части.

Заключение

Это конец первой части о внутренностях ядра Linux. В следующей части мы увидим первый код на языке C, который выполняется при настройке ядра Linux, реализацию процедур для работы с памятью, таких как `memset` , `memcpy` , `earlyprintk` , инициализацию консоли и многое другое.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

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Процесс загрузки ядра. Часть 2.

Первые шаги в настройке ядра

Мы начали изучение внутренностей Linux в предыдущей части и увидели начальную часть кода настройки ядра. Мы остановились на вызове функции `main` (это первая функция, написанная на C) из arch/x86/boot/main.c.

В этой части мы продолжим исследовать код установки ядра, а именно

- `защищённый режим`,
- переход в него,
- инициализацию кучи и консоли,
- обнаружение памяти, проверку CPU, инициализацию клавиатуры
- и многое другое.

Итак, давайте начнём.

Защищённый режим

Прежде чем мы сможем перейти к нативному для Intel64 режиму Long Mode, ядро должно переключить CPU в защищённый режим.

Что такое защищённый режим? Защищённый режим был впервые добавлен в архитектуре x86 в 1982 году и был основным режимом процессоров Intel, начиная с 80286, пока в Intel64 не появился режим Long Mode.

Основная причина не использовать режим реальных адресов заключается в том, что возможен лишь очень ограниченный доступ к оперативной памяти. Как вы помните из предыдущей части, есть только 2^{20} байт или 1 мегабайт, а иногда даже 640 килобайт оперативной памяти, доступной в режиме реальных адресов.

Защищённый режим принёс много изменений, но главным является отличие в управлении памятью. 20-битная адресная шина была заменена на 32-битную. Это позволило обеспечить доступ к 4 Гб памяти против 1 мегабайта в режиме реальных адресов. Также была добавлена поддержка страничной организации памяти, про которую вы можете прочитать в следующих разделах.

Управление памятью в защищённом режиме разделяется на две, почти независимые части:

- Сегментация
- Страничная организация

Здесь мы только о сегментации. Страничная организация будет обсуждаться в следующих разделах.

Как вы можете помнить из предыдущей части, адреса в режиме реальных адресов состоят из двух частей:

- Базовый адрес сегмента
- Смещение от базового сегмента

И мы можем получить физический адрес, если нам известны эти две части:

```
Физический адрес = Селектор сегмента * 16 + Смещение
```

Сегментация памяти в защищённом режиме была полностью переделана. В нём нет фиксированных 64 килобайтных сегментов. Вместо этого, размер и расположение каждого сегмента описывается структурой данных, называемой *дескриптором сегмента*. Дескрипторы сегментов хранятся в структуре данных под названием `глобальная дескрипторная` `таблица` (GDT).

GDT представляет собой структуру, которая находится в памяти. Она не имеет постоянного места в памяти, поэтому её адрес хранится в специальном регистре `GDTR`. Позже мы увидим загрузку GDT в коде ядра Linux. Там будет операция для её загрузки в память, что-то вроде:

```
lgdt gdt
```

где инструкция `lgdt` загружает базовый адрес и предел (размер) глобальной дескрипторной таблицы в регистр `GDTR`. `GDTR` является 48-битным регистром и состоит из двух частей:

- размер (16 бит) глобальной дескрипторной таблицы;
- адрес (32 бита) глобальной дескрипторной таблицы.

Как упоминалось ранее, GDT содержит `дескрипторы сегментов`, которые описывают сегменты памяти. Каждый дескриптор является 64-битным. Общая схема дескриптора такова:

```
 63                  56         51    48    45           39                 32
------------------------------------------------------------------------------
|                       | |B| |A|       | |   | |0|E|W|A|                     |
| БАЗОВЫЙ АДРЕС 31:24   |G|/|L|V| ПРЕДЕЛ |P|DPL|S|  ТИП  | БАЗОВЫЙ АДРЕС 23:16 |
|                       | |D| |L| 19:16  | |   | |1|C|R|A|                     |
------------------------------------------------------------------------------

 31                                  16 15                                   0
------------------------------------------------------------------------------
|                                        |                                   |
|          БАЗОВЫЙ АДРЕС 15:0            |             ПРЕДЕЛ 15:0           |
|                                        |                                   |
------------------------------------------------------------------------------
```

Не волнуйтесь, я знаю, после режима реальных адресов это выглядит немного страшно, но на самом деле всё довольно легко. Например, ПРЕДЕЛ 15:0 означает, что биты 0-15 предела расположены в начале дескриптора. Остальная часть находится в ПРЕДЕЛ 19:16, который расположен в битах 48-51 дескриптора. Таким образом, размер предела составляет 0-19, т.е 20 бит. Давайте внимательно взглянем на структуру дескриптора:

1. Предел (20 бит) находится в пределах 0-15, 48-51 бит. Он определяет `длину_сегмента - 1`. Зависит от бита `G` (гранулярность).

   - Если `G` (бит 55) и предел сегмента равен 0, то размер сегмента составляет 1 байт
   - Если `G` равен 1, а предел сегмента равен 0, то размер сегмента составляет 4096 байт
   - Если `G` равен 0, а предел сегмента равен 0xfffff, то размер сегмента составляет 1 мегабайт
   - Если `G` равен 1, а предел сегмента равен 0xfffff, то размер сегмента составляет 4 гигабайта

   Таким образом, если

   - G равен 0, предел интерпретируется в терминах 1 байта, а максимальный размер сегмента может составлять 1 мегабайт.
   - G равен 1, предел интерпретируется в терминах 4096 байт = 4 килобайта = 1 страница, а максимальный размер сегмента может составлять 4 гигабайта. На самом деле, когда G равен 1, значение предела сдвигается на 12 бит влево. Таким образом, 20 бит + 12 бит = 32 бита и 2^{32} = 4 гигабайта.
2. Базовый адрес (32 бита) находится в пределах 16-31, 32-39 и 56-63 бит. Он определяет физический адрес начального расположения сегмента.

3. Тип/Атрибут (5 бит) в пределах 40-47 бит определяет тип сегмента и виды доступа к нему.

   - Флаг `S` (бит 44) определяет тип дескриптора. Если `S` равен 0, то этот сегмент является системным сегментом, а если `S` равен 1, то этот сегмент является сегментом кода или сегментом данных (сегменты стека являются сегментами данных, которые должны быть сегментами для чтения/записи).

Для того чтобы определить, является ли сегмент сегментом кода или сегментом данных, мы можем проверить атрибут (бит 43), обозначенный как 0 в приведённой выше схеме. Если он равен 0, то сегмент является сегментом данных, в противном случае это сегмент кода.

Сегмент может быть одного из следующих типов:

```
|           Поле типа         | Тип дескриптора | Описание
|-----------------------------|-----------------|------------------
| Десятичное                  |                 |
|             0    E    W   A |                 |
| 0           0    0    0   0 | Данные          | Только для чтения
| 1           0    0    0   1 | Данные          | Только для чтения, было обращение
| 2           0    0    1   0 | Данные          | Чтение/запись
| 3           0    0    1   1 | Данные          | Чтение/запись, было обращение
| 4           0    1    0   0 | Данные          | Только для чтения, растёт вниз
| 5           0    1    0   1 | Данные          | Только для чтения, растёт вниз, было обращение
| 6           0    1    1   0 | Данные          | Чтение/запись, растёт вниз
| 7           0    1    1   1 | Данные          | Чтение/запись, растёт вниз, было обращение
|                  C    R   A |                 |
| 8           1    0    0   0 | Код             | Только для выполнения
| 9           1    0    0   1 | Код             | Только для выполнения, было обращение
| 10          1    0    1   0 | Код             | Выполнение/чтение
| 11          1    0    1   1 | Код             | Выполнение/чтение, было обращение
| 12          1    1    0   0 | Код             | Только для выполнения, подчинённый
| 14          1    1    0   1 | Код             | Только для выполнения, подчинённый, было обращение
| 13          1    1    1   0 | Код             | Выполнение/чтение, подчинённый
| 15          1    1    1   1 | Код             | Выполнение/чтение, подчинённый, было обращение
```

Как мы можем видеть, первый бит (бит 43) равен `0` для сегмента *данных* и `1` для сегмента *кода*. Следующие три бита (40, 41, 42): либо биты `EWA` (бит направления расширения (*E*xpansion), бит записи (*W*ritable), бит обращения (*A*ccessible)), либо `CRA` (бит подчинения (*C*onforming), бит чтения (*R*eadable), бит доступа (*A*ccessible)).

- Если E (бит 42) равен 0, то сегмент растёт вверх, в противном случае растёт вниз. Подробнее здесь.
- Если W (бит 41) (для сегмента данных) равен 1, то запись в сегмент разрешена. Обратите внимание, что право на чтение всегда разрешено для сегментов данных.
- A (бит 40) - было ли обращение процессора к сегменту.
- C (бит 43) - бит подчинения (для сегмента кода). Если C равен 1, то сегмент кода может быть выполнен из более низкого уровня привилегий, например, из уровня пользователя. Если C равно 0, то сегмент может быть выполнен только из того же уровня привилегий.
- R (бит 41) контролирует доступ чтения сегментам кода; когда он равен 1, то чтение сегмента разрешено. Право на запись всегда запрещено для сегмента кода.

1. DPL [2 бита] (уровень привилегий сегмента (Descriptor Privilege Level)) находится в 45-46 битах. Определяет уровень привилегий сегмента от 0 до 3, где 0 является самым привилегированным.

2. Флаг P (бит 47) - указывает на присутствие сегмента в памяти. Если P равен 0, то сегмент является *недействительным* и процессор откажется читать этот сегмент.

3. Флаг AVL (бит 52) - доступный и зарезервированный бит. Игнорируется в Linux.

4. Флаг L (бит 53) - указывает на то, содержит ли сегмент кода нативный 64-битный код. Если он равен 1, то сегмент кода будет выполнен в 64-битном режиме.

5. Флаг D/B (бит 54) - флаг разрядности (Default/Big), определяет размер операнда, т.е 16/32 бит. Если он установлен, то находящиеся в сегменте операнды считаются имеющими размер 32 бита, иначе 16 бит.

Сегментные регистры содержат селекторы сегментов, так же как и в режиме реальных адресов. Тем не менее, в защищённом режиме селектор сегмента обрабатывается иначе. Каждый дескриптор сегмента имеет соответствующий селектор сегмента, который представляет собой 16-битную структуру:

```
 15              3 2  1     0
-----------------------------
|      Index     | TI | RPL |
-----------------------------
```

Где,

* **Index** определяет номер дескриптора в GDT.
* **TI** (указатель таблицы (Table Indicator)) определяет таблицу, в которой нужно искать дескриптор. Если он равен 0, то поиск происходит в глобальной дескрипторной таблице (GDT), в противном случае в локальной дескрипторной таблице (LDT).
* **RPL** определяет уровень привилегий.

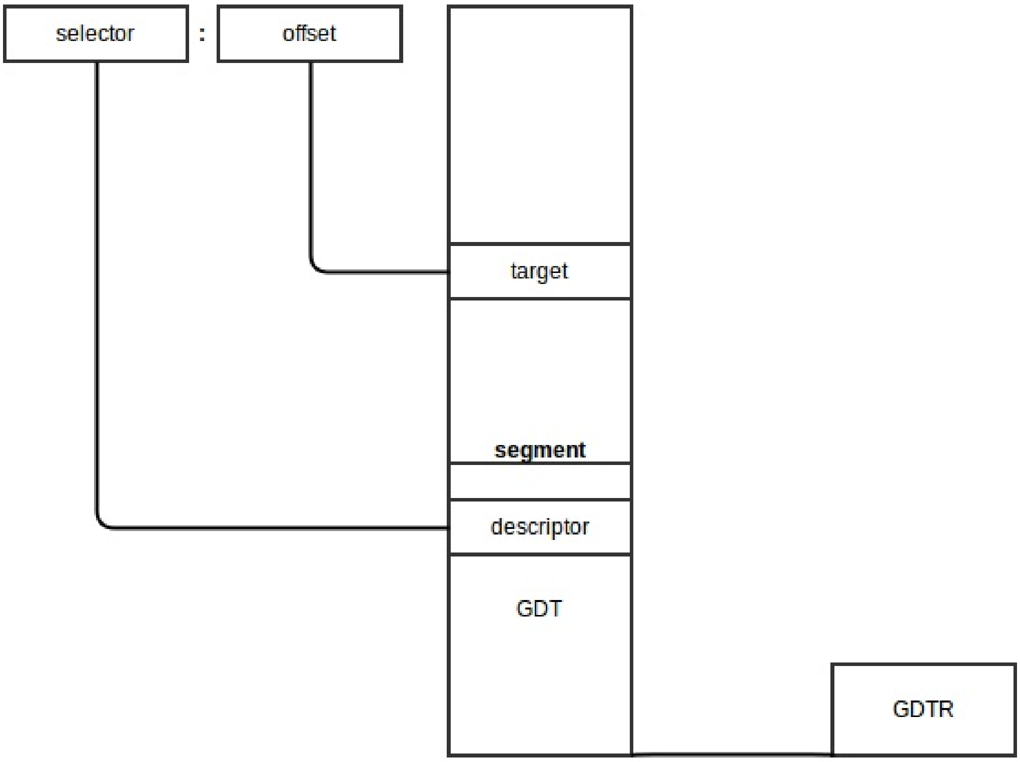
Каждый сегментный регистр имеет видимую и скрытую часть.

* Видимая - здесь хранится селектор сегмента.
* Скрытая - здесь хранится дескриптор сегмента, который содержит базовый адрес, предел, атрибуты, флаги.

Необходимы следующие шаги, чтобы получить физический адрес в защищённом режиме:

* Селектор сегмента должен быть загружен в один из сегментных регистров
* CPU пытается найти дескриптор сегмента по адресу `GDT + Index` из селектора и загрузить дескриптор в *скрытую* часть сегментного регистра
* Если страничная организация памяти отключена, линейный адрес сегмента или его физический адрес задается формулой: Базовый адрес (найденный в дескрипторе, полученном на предыдущем шаге) + Смещение.

Схематично это будет выглядеть следующим образом:



Алгоритм перехода из режима реальных адресов в защищённый режим:

- Отключить прерывания
- Описать и загрузить GDT инструкцией `lgdt`
- Установить бит PE (Protection Enable) в CR0 (регистр управления 0 (Control Register 0))
- Перейти к коду защищённого режима

Полный переход в защищённый режим мы увидим в следующей части, но прежде чем мы сможем перейти в защищённый режим, нужно совершить ещё несколько приготовлений.

Давайте посмотрим на arch/x86/boot/main.c. Мы можем видеть некоторые подпрограммы, которые выполняют инициализацию клавиатуры, инициализацию кучи и т.д. Рассмотрим их.

Копирование параметров загрузки в "нулевую страницу" (zeropage)

Мы стартуем из подпрограммы `main` в "main.c". Первая функция, которая вызывается в `main` - `copy_boot_params(void)`. Она копирует заголовок настройки ядра в поле структуры `boot_params`, которая определена в arch/x86/include/uapi/asm/bootparam.h.

Структура `boot_params` содержит поле `struct setup_header hdr`. Эта структура содержит те же поля, что и в протоколе загрузки Linux и заполняется загрузчиком, а так же во время компиляции/сборки ядра. `copy_boot_params` делает две вещи:

1. Копирует `hdr` из header.S в поле `setup_header` в структуре `boot_params`
2. Обновляет указатель на командную строку ядра, если ядро было загружено со старым протоколом командной строки.

Обратите внимание на то, что он копирует `hdr` с помощью функции `memcpy`, которая определена в copy.S. Взглянем на неё:

```
GLOBAL(memcpy)
    pushw   %si
    pushw   %di
    movw    %ax, %di
    movw    %dx, %si
    pushw   %cx
    shrw    $2, %cx
    rep; movsl
    popw    %cx
    andw    $3, %cx
    rep; movsb
    popw    %di
    popw    %si
    retl
ENDPROC(memcpy)
```

Да, мы только что перешли в C-код и снова вернулись к ассемблеру :) Прежде всего мы видим, что `memcpy` и другие подпрограммы, расположенные здесь, начинаются и заканчиваются двумя макросами: `GLOBAL` и `ENDPROC`. Макрос `GLOBAL` описан в arch/x86/include/asm/linkage.h и определяет директиву `globl`, а так же метку для него. `ENDPROC` описан в include/linux/linkage.h; отмечает символ `name` в качестве имени функции и заканчивается размером символа `name`.

Реализация `memcpy` достаточно проста. Во-первых, она помещает значения регистров `si` и `di` в стек для их сохранения, так как они будут меняться в течении работы. Как мы видим из `REALMODE_CFLAGS` в `arch/x86/Makefile` система сборки ядра использует параметр GCC `-mregparm = 3`, поэтому функции получают первые три параметра из регистров `ax`, `dx` и `cx`. Вызов `memcpy` выглядит следующим образом:

```
memcpy(&boot_params.hdr, &hdr, sizeof hdr);
```

Так,

- `ax` будет содержать адрес `boot_params.hdr`
- `dx` будет содержать адрес `hdr`
- `cx` будет содержать размер `hdr` в байтах.

`memcpy` помещает адрес `boot_params.hdr` в `di` и сохраняет размер в стеке. После этого она сдвигается вправо на 2 размера (или делит на 4) и копирует из `si` в `di` по 4 байта. Далее снова восстанавливает размер `hdr`, выравнивает по 4 байта и копирует остальную часть байтов из `si` в `di` побайтово (если они есть). В конце восстанавливает значения `si` и `di` из стека и после этого завершает копирование.

Инициализация консоли

После того как `hdr` скопирован в `boot_params.hdr`, следующим шагом является инициализация консоли с помощью вызова функции `console_init`, определённой в arch/x86/boot/early_serial_console.c.

Функция пытается найти опцию `earlyprintk` в командной строке и, если поиск завершился успехом, парсит адрес порта, скорость передачи данных и инициализирует последовательный порт. Значение опции `earlyprintk` может быть одним из следующих:

* serial,0x3f8,115200
* serial,ttyS0,115200
* ttyS0,115200

После инициализации последовательного порта мы можем увидеть первый вывод:

```
if (cmdline_find_option_bool("debug"))
    puts("early console in setup code\n");
```

Определение `puts` находится в tty.c. Как мы видим, она печатает символ за символом в цикле, вызывая функцию `putchar`. Давайте посмотрим на реализацию `putchar`:

```
void __attribute__((section(".inittext"))) putchar(int ch)
{
    if (ch == '\n')
        putchar('\r');

    bios_putchar(ch);

    if (early_serial_base != 0)
        serial_putchar(ch);
}
```

`__attribute__((section(".inittext")))` означает, что код будет находиться в секции `.inittext`. Мы можем найти его в файле компоновщика setup.ld.

Прежде всего, `putchar` проверяет наличие символа `\n` и, если он найден, печатает перед ним `\r`. После этого она выводит символ на экране VGA, вызвав BIOS с прерыванием `0x10`:

```
static void __attribute__((section(".inittext"))) bios_putchar(int ch)
{
    struct biosregs ireg;

    initregs(&ireg);
    ireg.bx = 0x0007;
    ireg.cx = 0x0001;
    ireg.ah = 0x0e;
    ireg.al = ch;
    intcall(0x10, &ireg, NULL);
}
```

`initregs` принимает структуру `biosregs` и в первую очередь заполняет `biosregs` нулями, используя функцию `memset`, а затем заполняет его значениями регистров:

```
    memset(reg, 0, sizeof *reg);
    reg->eflags |= X86_EFLAGS_CF;
    reg->ds = ds();
    reg->es = ds();
    reg->fs = fs();
    reg->gs = gs();
```

Давайте посмотри на реализацию memset:

```
GLOBAL(memset)
    pushw   %di
    movw    %ax, %di
    movzbl  %dl, %eax
    imull   $0x01010101,%eax
    pushw   %cx
    shrw    $2, %cx
```

```
    rep; stosl
    popw    %cx
    andw    $3, %cx
    rep; stosb
    popw    %di
    retl
ENDPROC(memset)
```

Как мы можем видеть, `memset` использует тоже самое соглашение о вызовах, как и `memcpy` : это означает, что функция получает свои параметры из регистров `ax` , `dx` и `cx` .

Как правило, реализация `memset` подобна реализации `memcpy` . Она сохраняет значение регистра `di` в стеке и помещает значение `ax` в `di` , которое является адресом структуры `biosregs` . Далее идёт инструкция `movzbl` , которая копирует значение `dl` в нижние 2 байта регистра `eax` . Оставшиеся 2 верхних байта `eax` будут заполнены нулями.

Следующая инструкция умножает `eax` на `0x01010101` . Это необходимо, так как `memset` будет копировать 4 байта одновременно. Например, нам нужно заполнить структуру, размер которой составляет 4 байта, значением `0x7` с помощью `memset` . В этом случае `eax` будет содержать значение `0x00000007` . Так что если мы умножим `eax` на `0x01010101` , мы получим `0x07070707` и теперь мы можем скопировать эти 4 байта в структуру. `memset` использует инструкцию `rep; stosl` для копирования `eax` в `es:di` .

Остальная часть `memset` делает почти то же самое, что и `memcpy` .

После того как структура `biosregs` заполнена с помощью `memset` , `bios_putchar` вызывает прерывание 0x10 для вывода символа. Затем она проверяет, инициализирован ли последовательный порт, и в случае если он инициализирован, записывает в него символ с помощью инструкций serial_putchar и `inb/outb` .

Инициализация кучи

После подготовки стека и BSS в header.S (смотрите предыдущую часть), ядро должно инициализировать кучу с помощью функции `init_heap` .

В первую очередь `init_heap` проверяет флаг `CAN_USE_HEAP` в `loadflags` в заголовке настройки ядра и если флаг был установлен, вычисляет конец стека:

```
    char *stack_end;

    if (boot_params.hdr.loadflags & CAN_USE_HEAP) {
        asm("leal %P1(%%esp),%0"
            : "=r" (stack_end) : "i" (-STACK_SIZE));
```

другими словами `stack_end = esp - STACK_SIZE` .

Затем идёт расчёт `heap_end` :

```
    heap_end = (char *)((size_t)boot_params.hdr.heap_end_ptr + 0x200);
```

что означает `heap_end_ptr` или `_end` + `512` (`0x200h`). Последняя проверка заключается в сравнении `heap_end` и `stack_end` . Если `heap_end` больше `stack_end` , то присваиваем `stack_end` значение `heap_end` , чтобы сделать их равными.

Теперь куча инициализирована и мы можем использовать её с помощью метода `GET_HEAP` . В следующих постах мы увидим как она используется, как её использовать и как она реализована.

Проверка CPU

Следующим шагом является проверка CPU с помощью функции `validate_cpu` из arch/x86/boot/cpu.c.

Она вызывает функцию `check_cpu` и передаёт ей два параметра: уровень CPU и необходимый уровень CPU; `check_cpu` проверяет, запущено ли ядро на нужном уровне CPU.

```
check_cpu(&cpu_level, &req_level, &err_flags);
if (cpu_level < req_level) {
    ...
    return -1;
}
```

`check_cpu` проверяет флаги CPU, наличие long mode в случае x86_64 (64-битного) CPU, проверяет поставщика процессора и делает специальные подготовки для некоторых производителей, такие как отключение SSE+SSE2 для AMD в случае их отсутствия и т.д.

На следующем этапе вы видим вызов функции `set_bios_mode`. Эта функция реализована только для режима `x86_64`:

```
static void set_bios_mode(void)
{
#ifdef CONFIG_X86_64
    struct biosregs ireg;

    initregs(&ireg);
    ireg.ax = 0xec00;
    ireg.bx = 2;
    intcall(0x15, &ireg, NULL);
#endif
}
```

Функция `set_bios_mode` выполняет прерывание `0x15`, чтобы сообщить BIOS, что будет использоваться long mode (если `bx == 2`).

Обнаружение памяти

Следующим шагом является обнаружение памяти с помощью функции `detect_memory`. `detect_memory` в основном предоставляет карту доступной оперативной памяти для CPU. Она использует различные программные интерфейсы для обнаружения памяти, такие как `0xe820`, `0xe801` и `0x88`. Здесь мы будем рассматривать только реализацию **0xE820**.

Давайте посмотрим на реализацию фуцнкции `detect_memory_e820` в arch/x86/boot/memory.c. Прежде всего, функция `detect_memory_e820` инициализирует структуру `biosregs`, как мы видели выше, и заполняет регистры специальными значениями для вызова `0xe820`:

```
    initregs(&ireg);
    ireg.ax  = 0xe820;
    ireg.cx  = sizeof buf;
    ireg.edx = SMAP;
    ireg.di  = (size_t)&buf;
```

- `ax` содержит номер функции (в нашем случае 0xe820)
- `cx` содержит размер буфера, который будет содержать данные о памяти
- `edx` должен содержать магическое число `SMAP`
- `es:di` должен содержать адрес буфера, который будет содержать данные из памяти
- `ebx` должен быть равен нулю.

Далее идёт цикл, в котором будут собраны данные о памяти. Он начинается с вызова BIOS прерывания `0x15`, который записывает одну строку из таблицы распределения адресов. Для получения следующей строки мы должны снова вызвать это прерывание (что мы и делаем в цикле). До следующего вызова `ebx` должен содержать значение, возвращённое ранее:

```
    intcall(0x15, &ireg, &oreg);
    ireg.ebx = oreg.ebx;
```

В конечном счёте мы делаем итерации в цикле для сбора данных из таблицы распределения адресов и записываем эти данные в массив `e820entry`:

- начало сегмента памяти
- размер сегмента памяти
- тип сегмента памяти (может ли конкретный сегмент быть использован или он зарезервирован).

Вы можете увидеть результат в выводе `dmesg`, что-то вроде:

```
[    0.000000] e820: BIOS-provided physical RAM map:
[    0.000000] BIOS-e820: [mem 0x0000000000000000-0x000000000009fbff] usable
[    0.000000] BIOS-e820: [mem 0x000000000009fc00-0x000000000009ffff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000000f0000-0x00000000000fffff] reserved
[    0.000000] BIOS-e820: [mem 0x0000000000100000-0x000000003ffdffff] usable
[    0.000000] BIOS-e820: [mem 0x000000003ffe0000-0x000000003fffffff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000fffc0000-0x00000000ffffffff] reserved
```

Инициализация клавиатуры

Следующим шагом является инициализация клавиатуры с помощью вызова функции `keyboard_init()`. Вначале `keyboard_init` инициализирует регистры с помощью функции `initregs`. Затем он вызывает прерывание 0x16 для получения статуса клавиатуры.

```
    initregs(&ireg);
    ireg.ah = 0x02;     /* Получение статуса клавиатуры */
    intcall(0x16, &ireg, &oreg);
    boot_params.kbd_status = oreg.al;
```

После этого она ещё раз вызывает 0x16 для установки частоты повторения и задержки.

```
    ireg.ax = 0x0305;   /* Установка частоты повторения клавиатуры */
    intcall(0x16, &ireg, NULL);
```

Выполнение запросов

Следующие несколько шагов - запросы для различных параметров. Мы не будем погружаться в подробности этих запросов, но вернёмся к этому в последующих частях. Давайте коротко взглянем на эти функции:

Первым шагом является получение информации Intel SpeedStep с помощью вызова функции `query_ist`. Она проверяет уровень CPU, и если он верный, вызывает прерывание `0x15` для получения информации и сохраняет результат в `boot_params`.

Следующая функция - query_apm_bios получает из BIOS информацию об Advanced Power Management. `query_apm_bios` также вызывает BIOS прерывание `0x15`, но с `ah = 0x53` для проверки поддержки `APM`. После выполнения `0x15`, функция `query_apm_bios` проверяет сигнатуру `PM` (она должна быть равна `0x504d`), флаг переноса (он должен быть равен 0, если есть поддержка `APM`) и значение регистра `cx` (оно должно быть равным 0x02, если имеется поддержка защищённого режима).

Далее она снова вызывает `0x15` , но с `ax = 0x5304` для отсоединения от интерфейса `APM` и подключению к интерфейсу 32-битного защищённого режима. В итоге она заполняет `boot_params.apm_bios_info` значениями, полученными из BIOS.

Обратите внимание: `query_apm_bios` будет выполняться только в том случае, если в конфигурационном файле установлен флаг времени компиляции `CONFIG_APM` или `CONFIG_APM_MODULE` :

```
#if defined(CONFIG_APM) || defined(CONFIG_APM_MODULE)
    query_apm_bios();
#endif
```

Последняя функция - `query_edd` , запрашивает из BIOS информацию об `Enhanced Disk Drive` . Давайте взглянем на реализацию `query_edd` .

В первую очередь она читает опцию edd из командной строки ядра и если она установлена в `off` , то `query_edd` завершает свою работу.

Если EDD включён, `query_edd` сканирует поддерживаемые BIOS жёсткие диски и запрашивает информацию о EDD в следующем цикле:

```
for (devno = 0x80; devno < 0x80+EDD_MBR_SIG_MAX; devno++) {
    if (!get_edd_info(devno, &ei) && boot_params.eddbuf_entries < EDDMAXNR) {
        memcpy(edp, &ei, sizeof ei);
        edp++;
        boot_params.eddbuf_entries++;
    }
    ...
    ...
    ...
}
```

где `0x80` - первый жёсткий диск, а значение макроса `EDD_MBR_SIG_MAX` равно 16. Она собирает данные в массив структур edd_info. `get_edd_info` проверяет наличие EDD путём вызова прерывания `0x13` с `ah = 0x41` и если EDD присутствует, `get_edd_info` снова вызывает `0x13` , но с `ah = 0x48` и `si` , содержащим адрес буфера, где будет храниться информация о EDD.

Заключение

Это конец второй части о внутренностях ядра Linux. В следующей части мы увидим настройки режима видео и остальные подготовки перед переходом в защищённый режим и непосредственно переход в него.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- Защищённый режим (Википедия)
- Защищённый режим (OSDEV)
- Long mode
- Неплохое объяснение режимов CPU с кодом

* Как использовать сегменты с ростом вниз на CPU Intel 386 и более поздних
* Документация по earlyprintk
* Параметры ядра
* Последовательная консоль
* Intel SpeedStep
* APM
* Спецификация EDD
* Документация TLDP для процесса загрузки Linux (старая)
* Предыдущая часть

Процесс загрузки ядра. Часть 3.

Инициализация видеорежима и переход в защищённый режим

Это третья часть серии `Процесса загрузки ядра`. В предыдущей части мы остановились прямо перед вызовом функции `set_video` из main.c. В этой части мы увидим:

- инициализацию видеорежима в коде настройки ядра,
- подготовку, сделанную перед переключением в защищённый режим,
- переход в защищённый режим

ПРИМЕЧАНИЕ: если вы ничего не знаете о защищённом режиме, вы можете найти некоторую информацию о нём в предыдущей части. Также есть несколько ссылок, которые могут вам помочь.

Как я уже писал ранее, мы начнём с функции `set_video`, которая определена в arch/x86/boot/video.c. Как мы можем видеть, она начинает работу с получения видеорежима из структуры `boot_params.hdr`:

```
u16 mode = boot_params.hdr.vid_mode;
```

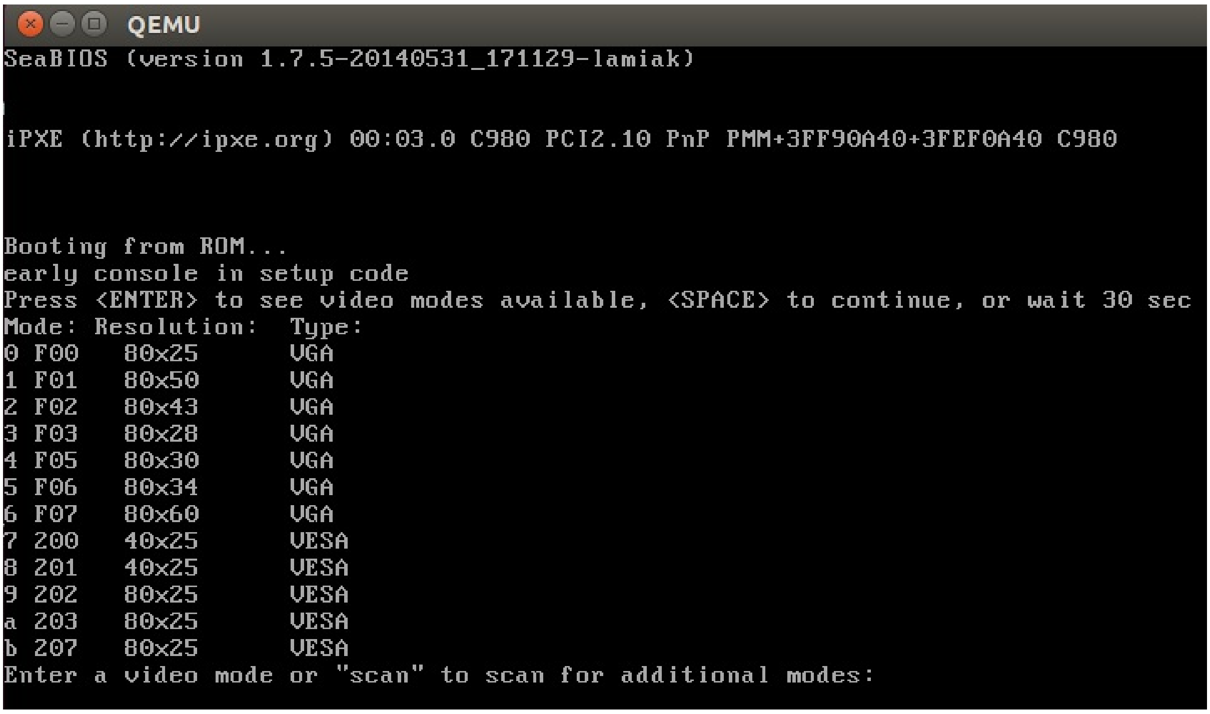
которую мы заполнили в функции `copy_boot_params` (вы можете прочитать об этом в предыдущем посте). `vid_mode` является обязательным полем, которое заполняется загрузчиком. Вы можете найти информацию об этом в протоколе загрузки ядра:

```
Offset    Proto    Name        Meaning
/Size
01FA/2    ALL        vid_mode    Video mode control
```

Как мы можем прочесть из протокола загрузки ядра Linux:

```
vga=<mode>
    <mode> может быть либо целочисленным значением (в C-нотации,
    десятичной, восьмеричной или шестнадцатеричной), либо одной из строк:
    "normal" (означает 0xFFFF), "ext" (означает 0xFFFE) или "ask"
    (означает 0xFFFD).  Это значение должно быть введено в поле
    vid_mode field, так как оно используется ядром до
    парсинга командной строки.
```

Таким образом, мы можем добавить параметр `vga` в конфигурационный файл GRUB (или любого другого загрузчика) и он передаст его в командную строку ядра. Как говорится в описании, этот параметр может иметь разные значения. Например, это может быть целым числом `0xFFFD` или `ask`. Если передать `ask` в `vga`, вы увидите примерно такое меню:



которое попросит выбрать видеорежим. Мы посмотрим на его реализацию, но перед этим рассмотрим некоторые другие вещи.

Типы данных ядра

Ранее мы видели определения различных типов данных в коде настройки ядра, таких как `u16` и т.д. Давайте взглянем на несколько типов данных, предоставляемых ядром:

Тип	char	short	int	long	u8	u16	u32	u64
Размер	1	2	4	8	1	2	4	8

Во время чтения исходного кода ядра вы будете часто встречать эти типы, так что было бы неплохо запомнить их.

API кучи

После того как мы получим `vid_mode` из `boot_params.hdr` в функции `set_video`, мы можем видеть вызов `RESET_HEAP`. `RESET_HEAP` представляет собой макрос, определённый в boot.h:

```
#define RESET_HEAP() ((void *)( HEAP = _end ))
```

Если вы читали вторую часть, то помните, что мы инициализировали кучу с помощью функции `init_heap`. У нас есть несколько полезных функций для кучи, которые определены в `boot.h`:

```
#define RESET_HEAP()
```

Как мы видели чуть выше, он сбрасывает кучу, установив переменную `HEAP` в `_end`, где `_end` просто `extern char _end[];`

Следующий макрос - `GET_HEAP`:

```
#define GET_HEAP(type, n) \
    ((type *)__get_heap(sizeof(type),__alignof__(type),(n)))
```

предназначен для выделения кучи. Он вызывает внутреннюю функцию `__get_heap` с тремя параметрами:

- размер типа данных, который должен быть выделен
- `__alignof__(type)` определяет, как переменные этого типа должны быть выровнены
- `n` определяет сколько элементов нужно выделить

Реализация `__get_heap` :

```
static inline char *__get_heap(size_t s, size_t a, size_t n)
{
    char *tmp;

    HEAP = (char *)(((size_t)HEAP+(a-1)) & ~(a-1));
    tmp = HEAP;
    HEAP += s*n;
    return tmp;
}
```

В дальнейшем мы увидим её использование, что-то вроде:

```
saved.data = GET_HEAP(u16, saved.x * saved.y);
```

Давайте попробуем понять принцип работы `__get_heap` . Мы видим, что `HEAP` (который равен `_end` после `RESET_HEAP()`) является адресом выровненной памяти в соответствии с параметром `a` . После этого мы сохраняем адрес памяти `HEAP` в переменную `tmp` , перемещаем `HEAP` в конец выделенного блока и возвращаем `tmp` , которая является начальным адресом выделенной памяти.

И последняя функция:

```
static inline bool heap_free(size_t n)
{
    return (int)(heap_end - HEAP) >= (int)n;
}
```

которая вычитает значение указателя `HEAP` из `heap_end` (мы вычисляли это в предыдущей части) и возвращает 1, если имеется достаточно памяти для `n` .

На этом всё. Теперь у нас есть простой API для кучи и можем перейти к настройке видеорежима.

Настройка видеорежима

Теперь мы можем перейти непосредственно к инициализации видеорежима. Мы остановились на вызове `RESET_HEAP()` в функции `set_video` . Далее идёт вызов функции `store_mode_params` , которая сохраняет параметры видеорежима в структуре `boot_params.screen_info` , определённой в include/uapi/linux/screen_info.h.

Если мы посмотрим на функцию `store_mode_params` , то увидим, что она начинается с вызова `store_cursor_position` . Как вы можете понять из названия функции, она получает информацию о курсоре и сохраняет её.

В первую очередь `store_cursor_position` инициализирует две переменные, которые имеют тип `biosregs` с `AH = 0x3` , и вызывает BIOS прерывание `0x10` . После того как прерывание успешно выполнено, она возвращает строку и столбец в регистрах `DL` и `DH` . Строка и столбец будут сохранены в полях `orig_x` и `orig_y` структуры `boot_params.screen_info` .

После выполнения `store_cursor_position` вызывается функция `store_video_mode`. Она просто получает текущий видеорежим и сохраняет его в `boot_params.screen_info.orig_video_mode`.

Далее она проверяет текущий видеорежим и устанавливает `video_segment`. После того как BIOS передаёт контроль в загрузочный сектор, для видеопамяти выделяются следующие адреса:

```
0xB000:0x0000     32 Кб   Видеопамять для монохромного текста
0xB800:0x0000     32 Кб   Видеопамять для цветного текста
```

Таким образом, мы устанавливаем переменную `video_segment` в `0xb000`, если текущий видеорежим MDA, HGC, или VGA в монохромном режиме, и в `0xb800`, если текущий видеорежим цветной. После настройки адреса видеофрагмента, размер шрифта должен быть сохранён в `boot_params.screen_info.orig_video_points`:

```
set_fs(0);
font_size = rdfs16(0x485);
boot_params.screen_info.orig_video_points = font_size;
```

В первую очередь мы устанавливаем регистр `FS` в 0 с помощью функции `set_fs`. В предыдущей части мы уже видели такие функции, как `set_fs`. Все они определены в boot.h. Далее мы читаем значение, которое находится по адресу `0x485` (эта область памяти используется для получения размера шрифта) и сохраняет размер шрифта в `boot_params.screen_info.orig_video_points`.

```
 x = rdfs16(0x44a);
 y = (adapter == ADAPTER_CGA) ? 25 : rdfs8(0x484)+1;
```

Далее мы получаем количество столбцов по адресу `0x44a`, и строк по адресу `0x484` и сохраняем их в `boot_params.screen_info.orig_video_cols` и `boot_params.screen_info.orig_video_lines`. После этого выполнение `store_mode_params` завершается.

Далее мы видим функцию `save_screen`, которая просто сохраняет содержимое экрана в куче. Эта функция собирает все данные, которые мы получили в предыдущей функции, такие как количество строк и столбцов и т.д, и сохраняет их в структуре `saved_screen`:

```
static struct saved_screen {
    int x, y;
    int curx, cury;
    u16 *data;
} saved;
```

Затем она проверяет, есть ли свободное место в куче:

```
if (!heap_free(saved.x*saved.y*sizeof(u16)+512))
        return;
```

и если места в куче достаточно, выделяет его и сохраняет в нём `saved_screen`.

Следующий вызов - `probe_cards(0)` из arch/x86/boot/video-mode.c. Она проходит по всем `video_cards` и собирает количество режимов, предоставляемых картой. И вот здесь интересный момент. Мы можем видеть цикл:

```
for (card = video_cards; card < video_cards_end; card++) {
  /* Здесь собираем количество режимов */
}
```

но `video_cards` нигде не объявлен. Ответ прост: каждый видеорежим, представленный в x86-коде настройки ядра, определён следующим образом:

```
static __videocard video_vga = {
    .card_name    = "VGA",
    .probe        = vga_probe,
    .set_mode    = vga_set_mode,
};
```

где `__videocard` - макрос:

```
#define __videocard struct card_info __attribute__((used,section(".videocards")))
```

который определяет структуру `card_info` :

```
struct card_info {
    const char *card_name;
    int (*set_mode)(struct mode_info *mode);
    int (*probe)(void);
    struct mode_info *modes;
    int nmodes;
    int unsafe;
    u16 xmode_first;
    u16 xmode_n;
};
```

которая находится в сегменте `.videocards` . Давайте посмотрим на скрипт компоновщика arch/x86/boot/setup.ld, в котором мы можем найти:

```
    .videocards    : {
        video_cards = .;
        *(.videocards)
        video_cards_end = .;
    }
```

Это значит, что `video_cards` - просто адрес в памяти, и все структуры `card_info` размещаются в этом сегменте. Это также означает, что все структуры `card_info` размещаются между `video_cards` и `video_cards_end` , и мы можем воспользоваться этим, чтобы пройтись по ним в цикле. После выполнения `probe_cards` у нас есть все структуры `static __videocard video_vga` с заполненными `nmodes` (число видеорежимов).

После завершения выполнения `probe_cards` , мы переходим в главный цикл функции `set_video` . Это бесконечный цикл, который пытается установить видеорежим с помощью функции `set_mode` и выводит меню, если установлен флаг `vid_mode=ask` командной строки ядра или если видеорежим не определён.

Функция `set_mode` определена в video-mode.c и принимает только один параметр - `mode` , который определяет количество видеорежимов (мы получили его из меню или в начале `setup_video` , из заголовка настройки ядра).

`set_mode` проверяет `mode` и вызывает функцию `raw_set_mode` . `raw_set_mode` вызывает `set_mode` для выбранной карты, т.е. `card->set_mode(struct mode_info*)` . Мы можем получить доступ к этой функции из структуры `card_info` . Каждый видеорежим определяет эту структуру со значениями, заполненными в зависимости от режима видео (например, для `vga` это функция `video_vga.set_mode` . См. выше пример структуры `card_info` для `vga`). `video_vga.set_mode` является `vga_set_mode` , который проверяет VGA-режим и вызывает соответствующую функцию:

```
static int vga_set_mode(struct mode_info *mode)
{
    vga_set_basic_mode();

    force_x = mode->x;
    force_y = mode->y;

    switch (mode->mode) {
```

```
    case VIDEO_80x25:
        break;
    case VIDEO_8POINT:
        vga_set_8font();
        break;
    case VIDEO_80x43:
        vga_set_80x43();
        break;
    case VIDEO_80x28:
        vga_set_14font();
        break;
    case VIDEO_80x30:
        vga_set_80x30();
        break;
    case VIDEO_80x34:
        vga_set_80x34();
        break;
    case VIDEO_80x60:
        vga_set_80x60();
        break;
    }
    return 0;
}
```

Каждая функция, которая устанавливает видеорежим, просто вызывает BIOS прерывание `0x10` с определённым значением в регистре `AH`.

После того как мы установили видеорежим, мы передаём его в `boot_params.hdr.vid_mode`.

Далее вызывается `vesa_store_edid`. Эта функция сохраняет информацию о EDID (**E**xtended **D**isplay **I**dentification **D**ata) для использования ядром. После этого снова вызывается `store_mode_params`. И наконец, если установлен `do_restore`, экран восстанавливается в предыдущее состояние.

Сделав это, мы завершаем настройку видеорежима и мы можем переключится в защищённый режим.

Последняя подготовка перед переходом в защищённый режим

Мы можем видеть последний вызов функции - `go_to_protected_mode` - в main.c. Как говорится в комментарии: `Do the last things and invoke protected mode`, так что давайте посмотрим на эти последние вещи и перейдём в защищённый режим.

Функция `go_to_protected_mode` определена в arch/x86/boot/pm.c. Она содержит функции, которые совершают последние приготовления, прежде чем мы сможем перейти в защищённый режим, так что давайте посмотрим на них и попытаться понять, что они делают и как это работает.

Во-первых, это вызов функции `realmode_switch_hook` в `go_to_protected_mode`. Эта функция вызывает хук переключения режима реальных адресов, если он присутствует, и выключает NMI. Хуки используются, если загрузчик работает во "враждебной" среде. Вы можете прочитать больше о хуках в протоколе загрузки (см. **ADVANCED BOOT LOADER HOOKS**).

Хук `realmode_switch` представляет собой указатель на 16-битную удалённую подпрограмму режима реальных адресов, которая отключает немаскируемые прерывания. После проверки хука `realmode_switch`, происходит выключение Non-Maskable Interrupts (NMI):

```
asm volatile("cli");
outb(0x80, 0x70);    /* Выключение NMI */
io_delay();
```

Первой вызывается ассемблерная инструкция `cli`, которая очищает флаг прерывания (`IF`). После этого внешние прерывания отключены. Следующая строка отключает NMI (немаскируемые прерывания).

Прерывание является сигналом, который отправляется CPU от аппаратного или программного обеспечения. После получения сигнала, CPU приостанавливает текущую последовательность команд, сохраняет своё состояние и передаёт управление обработчику прерываний. После того как обработчик прерывания закончил свою работу, он передаёт управление прерванной инструкции. Немаскируемые прерывания (NMI) - это прерывания, которые обрабатываются всегда, независимо от запретов на другие прерывания. Их нельзя игнорировать, и, как правило, они используются для подачи сигнала о невосстанавливаемых аппаратных ошибок. Сейчас мы не будем погружаться в детали прерываний, но обсудим это в следующих постах.

Давайте вернёмся к коду. Мы видим, что вторая строка пишет байт `0x80` (отключённый бит) в `0x70` (регистр CMOS Address). После этого происходит вызов функции `io_delay`. `io_delay` вызывает небольшую задержку и выглядит следующим образом:

```
static inline void io_delay(void)
{
    const u16 DELAY_PORT = 0x80;
    asm volatile("outb %%al,%0" : : "dN" (DELAY_PORT));
}
```

Для вывода любого байта в порт `0x80` необходима задержка в 1 мкс. Таким образом, мы можем записать любое значение (в нашем случае значение из регистра `AL`) в порт `0x80`. После задержки, функция `realmode_switch_hook` завершает выполнение и мы можем перейти к следующей функции.

Следующая функция - `enable_a20` - включает линию A20. Она определена в arch/x86/boot/a20.c и совершает попытку включения шлюза адресной линии A20 различными методами. Первым из них является функция `a20_test_short`, которая проверят, является ли A20 включённой или нет с помощью функции `a20_test`:

```
static int a20_test(int loops)
{
    int ok = 0;
    int saved, ctr;

    set_fs(0x0000);
    set_gs(0xffff);

    saved = ctr = rdfs32(A20_TEST_ADDR);

    while (loops--) {
        wrfs32(++ctr, A20_TEST_ADDR);
        io_delay();    /* Serialize and make delay constant */
        ok = rdgs32(A20_TEST_ADDR+0x10) ^ ctr;
        if (ok)
            break;
    }

    wrfs32(saved, A20_TEST_ADDR);
    return ok;
}
```

В первую очередь мы устанавливаем регистр `FS` в `0x0000` и регистр `GS` в `0xffff`. Далее мы читаем значение по адресу `A20_TEST_ADDR` (`0x200`) и сохраняем его в переменную `saved` и `ctr`.

После этого мы записываем обновлённое значение `ctr` в `fs:A20_TEST_ADDR` или `fs:0x200` с помощью функции `wrfs32`, совершаем задержку в 1 мс, а затем читаем значение из регистра `GS` по адресу `A20_TEST_ADDR+0x10`. В случае, если линия A20 отключена, адрес будет перекрыт, в противном случае, если он не равен нулю, линия A20 уже включена.

Если линия A20 отключена, мы пытаемся включить её с помощью других методов, которые вы можете найти в `a20.c`. Например, это может быть сделано с помощью вызова BIOS прерывания `0x15` с `AH=0x2041` и т.д.

Если функция `enabled_a20` завершается неудачей, выводится сообщение об ошибке и вызывается функция `die`. Вы можете вспомнить её из первого файла исходного кода, откуда мы начали - arch/x86/boot/header.S:

```
die:
    hlt
    jmp    die
    .size    die, .-die
```

После того как шлюз линии A20 успешно включён, вызывается функция `reset_coprocessor`:

```
outb(0, 0xf0);
outb(0, 0xf1);
```

Она очищает математический сопроцессор путём записи `0` в `0xf0`, а затем сбрасывает его при помощи записи `0` в `0xf1`.

После этого вызывается функция `mask_all_interrupts`:

```
outb(0xff, 0xa1);       /* Маскирует все прерывания на вторичном PIC */
outb(0xfb, 0x21);       /* Маскирует все, кроме каскада на первичном PIC */
```

Она маскирует все прерывания на вторичном PIC (программируемый контроллер прерываний) и первичном PIC, за исключением IRQ2 на первичном PIC.

И теперь, после всех приготовлений, мы можем увидеть фактический переход в защищённый режим.

Настройка таблицы векторов прерываний

Теперь мы настраиваем таблицу векторов прерываний (IDT). Функция `setup_idt`:

```
static void setup_idt(void)
{
    static const struct gdt_ptr null_idt = {0, 0};
    asm volatile("lidtl %0" : : "m" (null_idt));
}
```

настраивает таблицу векторов прерываний (описывает обработчики прерываний и т.д). В настоящее время IDT не установлена (мы увидим это позже), сейчас мы просто загрузили IDT инструкцией `lidtl`. `null_idt` содержит адрес и размер IDT, но сейчас они равны нулю. `null_idt` является структурой `gdt_ptr` и определена следующим образом:

```
struct gdt_ptr {
    u16 len;
    u32 ptr;
} __attribute__((packed));
```

где мы можем видеть 16-битную длину (`len`) IDT и 32-битный указатель на неё (более подробно о IDT и прерываниях вы увидите в следующих постах). `__attribute__((packed))` означает, что размер `gdt_ptr` является минимальным требуемым размером. Таким образом, размер `gdt_ptr` должен быть равен 6 байтам или 48 битам. (Далее мы будем загружать указатель на `gdt_ptr` в регистр `GDTR` и вы, возможно, помните из предыдущего поста, что это 48-битный регистр).

Настройка глобальной таблицы дескрипторов

Далее идёт настройка глобальной таблицы дескрипторов (GDT). Мы можем видеть функцию `setup_gdt`, которая настраивает GDT (вы можете прочитать про это в посте Процесс загрузки ядра. Часть 2.). В этой функции определён массив `boot_gdt`, который содержит определение трёх сегментов:

```
    static const u64 boot_gdt[] __attribute__((aligned(16))) = {
        [GDT_ENTRY_BOOT_CS] = GDT_ENTRY(0xc09b, 0, 0xfffff),
        [GDT_ENTRY_BOOT_DS] = GDT_ENTRY(0xc093, 0, 0xfffff),
        [GDT_ENTRY_BOOT_TSS] = GDT_ENTRY(0x0089, 4096, 103),
    };
```

для получения кода, данных и TSS (Task State Segment, сегмент состояния задачи). В данный момент мы не будем использовать сегмент состояния задачи. Как мы можем видеть в строке комментария, он был добавлен специально для Intel VT (здесь вы можете найти коммит, который описывает его). Давайте посмотри на `boot_gdt`. Прежде всего отметим, что она имеет атрибут `__attribute__((aligned(16)))`. Это означает, что структура будет выровнена по 16 байтам. Давайте посмотрим на простой пример:

```
#include <stdio.h>

struct aligned {
    int a;
}__attribute__((aligned(16)));

struct nonaligned {
    int b;
};

int main(void)
{
    struct aligned    a;
    struct nonaligned na;

    printf("Not aligned - %zu \n", sizeof(na));
    printf("Aligned - %zu \n", sizeof(a));

    return 0;
}
```

Технически, структура, которая содержит одно поле типа `int`, должна иметь размер 4 байта, но для `aligned` структуры потребуется 16 байт для хранения в памяти:

```
$ gcc test.c -o test && test
Not aligned - 4
Aligned - 16
```

Здесь `GDT_ENTRY_BOOT_CS` имеет индекс - 2, `GDT_ENTRY_BOOT_DS` является `GDT_ENTRY_BOOT_CS + 1` и т.д. Он начинается с 2, поскольку первый является обязательным нулевым дескриптором (индекс - 0), а второй не используется (индекс - 1).

`GDT_ENTRY` - это макрос, который принимает флаги, базовый адрес, предел и создаёт запись в GDT. Для примера посмотрим на запись сегмента кода. `GDT_ENTRY` принимает следующие значения:

- базовый адрес - `0`
- предел - `0xfffff`
- флаги - `0xc09b`

Что это значит? Базовый адрес сегмента равен 0, а предел (размер сегмента) равен `0xffff` (1 Мб). Давайте посмотрим на флаги. В двоичном виде значение `0xc09b` будет выглядеть следующим образом:

```
1100 0000 1001 1011
```

Попробуем понять, что означает каждый бит. Мы пройдёмся по всем битам слева направо

- 1 - (G) бит гранулярности
- 1 - (D) если равен 0 - 16-битный сегмент; 1 - 32-битный сегмент
- 0 - (L) если 1 - выполняется в 64-битном режиме
- 0 - (AVL) доступен для использования системным ПО
- 0000 - 4 бита предела в 19:16 бит в дескрипторе
- 1 - (P) присутствие сегмента в памяти
- 00 - (DPL) - уровень привилегий, 0 является высшей привилегией
- 1 - (S) сегмент кода или данных, не системный сегмент
- 101 - тип сегмента и виды доступа к нему (чтение, выполнение)
- 1 - бит обращения

Вы можете прочитать больше о каждом бите в предыдущем посте или в документации для разработчиков ПО на архитектуре Intel® 64 и IA-32.

После этого мы получаем длину GDT:

```
gdt.len = sizeof(boot_gdt)-1;
```

Здесь мы получаем размер `boot_gdt` и вычитаем 1 (последний действительный адрес в GDT).

Далее получаем указатель на GDT:

```
gdt.ptr = (u32)&boot_gdt + (ds() << 4);
```

Здесь мы просто получаем адрес `boot_gdt` и добавляем его к адресу сегмента данных, сдвинутого влево на 4 бита (не забывайте, что сейчас мы находимся в режиме реальных адресов).

И наконец, мы выполняем инструкцию `lgdtl` для загрузки GDT в регистр GDTR:

```
asm volatile("lgdtl %0" : : "m" (gdt));
```

Фактический переход в защищённый режим

Это конец функции `go_to_protected_mode`. Мы загрузили IDT, GDT, отключили прерывания и теперь можем переключить CPU в защищённый режим. Последний шаг - вызов функции `protected_mode_jump` с двумя параметрами:

```
protected_mode_jump(boot_params.hdr.code32_start, (u32)&boot_params + (ds() << 4));
```

которая определена в arch/x86/boot/pmjump.S. Она получает два параметра:

- адрес точки входа в защищённый режим
- адрес `boot_params`

Давайте заглянем внутрь `protected_mode_jump`. Как я уже писал выше, вы можете найти его в `arch/x86/boot/pmjump.S`. Первый параметр находится в регистре `eax`, второй в `edx`.

В первую очередь мы помещаем адрес `boot_params` в регистр `esi` и адрес регистра сегмента кода `cs` в `bx`. Далее мы сдвигаем `bx` на 4 бита и добавляем к нему адрес метки `2` (`(cs << 4) + in_pm32`, физический адрес для "прыжка" после перехода в 32-битный режим) и переходим на метку `1`. После этого `in_pm32` в метке `2` будет перезаписан следующим образом: `(cs << 4) + in_pm32` .

Далее мы помещаем сегмент данных и сегмент состояния задачи в регистры `cx` и `di` :

```
movw    $__BOOT_DS, %cx
movw    $__BOOT_TSS, %di
```

Как вы можете прочесть выше, `GDT_ENTRY_BOOT_CS` имеет индекс 2 и каждая запись GDT имеет размер 8 байт, поэтому `CS` будет `2 * 8 = 16` , `__BOOT_DS` равен 24 и т.д.

Далее мы устанавливаем бит `PE` (Protection Enable) в регистре управления `CR0` :

```
movl    %cr0, %edx
orb    $X86_CR0_PE, %dl
movl    %edx, %cr0
```

и совершаем длинный переход в защищённый режим:

```
    .byte    0x66, 0xea
2:    .long    in_pm32
    .word    __BOOT_CS
```

где

- `0x66` - префикс размера операнда, который позволяет смешивать как 16-битный, так и 32-битный код,
- `0xea` - опкод инструкции перехода,
- `in_pm32` - смещение сегмента в защищённом режиме, которое имеет значение `(cs << 4) + in_pm32` , полученное из режима реальных адресов
- `__BOOT_CS` - сегмент кода, на который мы хотим перейти.

После этого мы наконец-то в защищённом режиме:

```
.code32
.section ".text32","ax"
```

Давайте посмотрим на первые шаги в защищённом режиме. Прежде всего мы устанавливаем сегмент данных следующим образом:

```
movl    %ecx, %ds
movl    %ecx, %es
movl    %ecx, %fs
movl    %ecx, %gs
movl    %ecx, %ss
```

Если вы обратили внимание, то можете вспомнить, что мы сохраняли `$__BOOT_DS` в регистре `cx` . Теперь мы заполнили все сегментные регистры, кроме `cs` (`cs` уже `__BOOT_CS`). Далее мы обнуляем все регистры общего назначения, кроме `eax` :

```
xorl    %ecx, %ecx
xorl    %edx, %edx
xorl    %ebx, %ebx
xorl    %ebp, %ebp
xorl    %edi, %edi
```

И в конце переходим к 32-битной точке входа:

```
jmpl    *%eax
```

Как вы помните, `eax` содержит адрес 32-битной записи (мы передали его как первый параметр в `protected_mode_jump`).

На этом всё. Теперь мы находимся в защищённом режиме и останавливаемся на этой точке входа. Что произойдёт дальше, мы увидим в следующей части.

Заключение

Это конец третьей части о внутренностях ядра Linux. В следующей части мы рассмотрим первые шаги в защищённом режиме и переход в long mode.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- VGA
- VESA BIOS Extensions
- Выравнивание данных
- Немаскируемое прерывание
- Линия A20
- GCC designated inits (назначенные инициализаторы)
- Атрибуты типов GCC
- Предыдущий пост

Процесс загрузки ядра. Часть 4.

Переход в 64-битный режим

Это четвёртая часть `Процесса загрузки ядра`, в которой вы увидите первые шаги в защищённом режиме, такие как проверка поддержки процессором long mode и SSE, страничная организация памяти, инициализация таблиц страниц и в конце мы обсудим переход в long mode.

ЗАМЕЧАНИЕ: данная часть содержит много ассемблерного кода, так что если вы не знакомы с ним, вы можете прочитать соответствующую литературу

В предыдущей части мы остановились на переходе к 32-битной точке входа в arch/x86/boot/pmjump.S:

```
jmpl    *%eax
```

Вы помните, что регистр `eax` содержит адрес 32-битной точки входа. Мы можем прочитать об этом в протоколе загрузки ядра Linux x86:

```
При использовании bzImage ядро в защищённом режиме перемещается на 0x100000
```

Давайте удостоверимся в том, что это правда, посмотрев на значения регистров в 32-битной точке входа:

```
eax            0x100000    1048576
ecx            0x0        0
edx            0x0        0
ebx            0x0        0
esp            0x1ff5c    0x1ff5c
ebp            0x0        0x0
esi            0x14470    83056
edi            0x0        0
eip            0x100000    0x100000
eflags         0x46        [ PF ZF ]
cs             0x10    16
ss             0x18    24
ds             0x18    24
es             0x18    24
fs             0x18    24
gs             0x18    24
```

Мы видим, что регистр `cs` содержит `0x10` (как вы помните из предыдущей части, это второй индекс в глобальной таблице дескрипторов), регистр `eip` содержит `0x100000`, и базовый адрес всех сегментов, в том числе сегмента кода, равен нулю. Таким образом, мы можем получить физический адрес - это будет `0:0x100000` или просто `0x100000`, как указано в протоколе загрузки. Давайте начнём с 32-битной точки входа.

32-битная точка входа

Мы можем найти определение 32-битной точки входа в arch/x86/boot/compressed/head_64.S:

```
    __HEAD
    .code32
ENTRY(startup_32)
....
....
....
```

```
ENDPROC(startup_32)
```

Прежде всего, почему директория `compressed`? На самом деле, `bzimage` является сжатым `vmlinux + заголовок + код настройки ядра`. Мы видели код настройки ядра во всех предыдущих частях. Таким образом, главная цель `head_64.S` - подготовка перехода в long mode, переход в него и декомпрессия ядра. В этой части мы увидим все шаги, вплоть до декомпрессии ядра.

В директории `arch/x86/boot/compressed` содержится два файла:

- head_32.S
- head_64.S

но мы будем рассматривать только `head_64.S`, потому что, как вы помните, эта книга только о `x86_64`; `head_32.S` в нашем случае не используется. Давайте посмотрим на arch/x86/boot/compressed/Makefile. Здесь мы можем увидить следующую цель сборки:

```
vmlinux-objs-y := $(obj)/vmlinux.lds $(obj)/head_$(BITS).o $(obj)/misc.o \
    $(obj)/string.o $(obj)/cmdline.o \
    $(obj)/piggy.o $(obj)/cpuflags.o
```

Обратите внимание на `$(obj)/head_$(BITS).o`. Это означает, что выбор файла (head_32.o или head_64.o) для линковки будет зависеть от значения `$(BITS)`. `$(BITS)` определён в arch/x86/Makefile, основанном на .config файле:

```
ifeq ($(CONFIG_X86_32),y)
        BITS := 32
        ...
        ...
else
        BITS := 64
        ...
        ...
endif
```

Перезагрузка сегментов, если это необходимо

Как было отмечено выше, мы начинаем с ассемблерного файла arch/x86/boot/compressed/head_64.S. Во-первых, мы видим определение специального атрибута секции перед определением `startup_32`:

```
    __HEAD
    .code32
ENTRY(startup_32)
```

`__HEAD` является макросом, определённым в include/linux/init.h и представляет собой следующую секцию:

```
#define __HEAD        .section    ".head.text","ax"
```

с именем `.head.text` и флагами `ax`. В нашем случае эти флаги означают, что секция является исполняемой или, другими словами, содержит код. Мы можем найти определение этой секции в скрипте компоновщика arch/x86/boot/compressed/vmlinux.lds.S:

```
SECTIONS
{
    . = 0;
    .head.text : {
        _head = . ;
        HEAD_TEXT
```

```
        _ehead = . ;
    }
    ...
    ...
    ...
}
```

Если вы не знакомы с синтаксисом скриптового языка компоновщика `GNU LD`, вы можете найти более подробную информацию в документации. Вкратце, символ `.` является специальной переменной компоновщика - счётчиком местоположения. Значение, присвоенное ему - это смещение по отношению к смещению сегмента. В нашем случае мы устанавливаем счётчик местоположения в ноль. Это означает, что наш код слинкован для запуска в памяти со смещения `0`. Кроме того, мы можем найти эту информацию в комментарии:

```
Be careful parts of head_64.S assume startup_32 is at address 0.
```

Хорошо, теперь мы знаем, где мы находимся, и сейчас самое время заглянуть внутрь функции `startup_32`.

В начале `startup_32` мы видим инструкцию `cld`, которая очищает бит `DF` в регистре флагов. Когда флаг направления очищен, все строковые операции, такие как stos, scas и др. будут инкрементировать индексные регистры `esi` или `edi`. Нам нужно очистить флаг направления, потому что позже мы будем использовать строковые операции для очистки пространства для таблиц страниц и т.д.

После того как бит `DF` очищен, следующим шагом является проверка флага `KEEP_SEGMENTS` из поля `loadflags` заголовка настройки ядра. Если вы помните, мы уже видели `loadflags` в самой первой части книги. Там мы проверяли флаг `CAN_USE_HEAP` чтобы узнать, можем ли мы использовать кучу. Теперь нам нужно проверить флаг `KEEP_SEGMENTS`. Данный флаг описан в протоколе загрузки:

```
Бит 6 (запись): KEEP_SEGMENTS
  Протокол: 2.07+
  - Если 0, перезагрузить регистры сегмента в 32-битной точке входа.
  - Если 1, не перезагружать регистры сегмента в 32-битной точке входа.
    Предполагается, что %cs %ds %ss %es установлены в плоские сегменты
    с базовым адресом 0 (или эквивалент для их среды).
```

Таким образом, если бит `KEEP_SEGMENTS` в `loadflags` не установлен, то сегментные регистры `ds`, `ss` и `es` должны быть установлены в индекс сегмента данных с базовым адресом `0`. Что мы и делаем:

```
    testb $KEEP_SEGMENTS, BP_loadflags(%esi)
    jnz 1f

    cli
    movl    $(__BOOT_DS), %eax
    movl    %eax, %ds
    movl    %eax, %es
    movl    %eax, %ss
```

Вы помните, что `__BOOT_DS` равен `0x18` (индекс сегмента данных в глобальной таблице дескрипторов). Если `KEEP_SEGMENTS` установлен, мы переходим на ближайшую метку `1f`, иначе обновляем сегментные регистры значением `__BOOT_DS`. Сделать это довольно легко, но есть один интересный момент. Если вы читали предыдущую часть, то помните, что мы уже обновили сегментные регистры сразу после перехода в защищённый режим в arch/x86/boot/pmjump.S. Так почему же нам снова нужно обновить значения в сегментных регистрах? Ответ прост. Ядро Linux также имеет 32-битный протокол загрузки и если загрузчик использует его для загрузки ядра, то весь код до `startup_32` будет пропущен. В этом случае `startup_32` будет первой точкой входа в ядро, и нет никаких гарантий, что сегментные регистры будут находиться в ожидаемом состоянии.

После того как мы проверили флаг `KEEP_SEGMENTS` и установили правильное значение в сегментные регистры, следующим шагом будет вычисление разницы между адресом, по которому мы загружены, и адресом, который был указан во время компиляции. Вы помните, что `setup.ld.S` содержит следующее определение в начале секции: `.head.text` : `. = 0` . Это значит, что код в этой секции скомпилирован для запуска по адресу `0` . Мы можем видеть это в выводе `objdump` :

```
arch/x86/boot/compressed/vmlinux:     file format elf64-x86-64


Disassembly of section .head.text:

0000000000000000 <startup_32>:
   0:   fc                      cld
   1:   f6 86 11 02 00 00 40    testb  $0x40,0x211(%rsi)
```

Утилита `objdump` говорит нам о том, что адрес `startup_32` равен `0` . Но на самом деле это не так. Наша текущая цель состоит в том, чтобы узнать настоящее местоположение. Довольно просто сделать это в long mode, поскольку он поддерживает относительную адресацию с помощью указателя `rip` , но в настоящее время мы находимся в защищённом режиме. Для того чтобы узнать адрес `startup_32` , мы будем использовать общепринятый шаблон. Нам необходимо определить метку, перейти на эту метку и вытолкнуть вершину стека в регистр:

```
call label
label: pop %reg
```

После этого регистр `%reg` будет содержать адрес метки. Давайте посмотрим на аналогичный код поиска адреса `startup_32` в ядре Linux:

```
        leal    (BP_scratch+4)(%esi), %esp
        call    1f
1:      popl    %ebp
        subl    $1b, %ebp
```

Как вы помните из предыдущей части, регистр `esi` содержит адрес структуры boot_params, которая была заполнена до перехода в защищённый режим. Структура `boot_params` содержит специальное поле `scratch` со смещением `0x1e4` . Это 4 байтное поле будет временным стеком для инструкции `call` . Мы получаем адрес поля `scratch` + `4` байта и помещаем его в регистр `esp` . Мы добавили `4` байта к базовому адресу поля `BP_scratch` , поскольку поле является временным стеком, а стек на архитектуре `x86_64` растёт сверху вниз. Таким образом, наш указатель стека будет указывать на вершину стека. Далее мы видим наш шаблон, который я описал ранее. Мы переходим на метку `1f` и помещаем её адрес в регистр `ebp` , потому что после выполнения инструкции `call` на вершине стека находится адрес возврата. Теперь у нас есть адрес метки `1f` и мы легко сможем получить адрес `startup_32` . Нам просто нужно вычесть адрес метки из адреса, который мы получили из стека:

```
startup_32 (0x0)       +-----------------------+
                       |                       |
                       |                       |
                       |                       |
                       |                       |
                       |                       |
                       |                       |
                       |                       |
                       |                       |
1f (смещение 0x0 + 1f) +-----------------------+ %ebp - реальный физический адрес
                       |                       |
                       |                       |
                       +-----------------------+
```

`startup_32` слинкован для запуска по адресу `0x0` и это значит, что `1f` имеет адрес `0x0 + смещение 1f`, примерно `0x21` байт. Регистр `ebp` содержит реальный физический адрес метки `1f`. Таким образом, если вычесть `1f` из `ebp`, мы получим реальный физический адрес `startup_32`. В протоколе загрузки ядра Linux описано, что базовый адрес ядра в защищённом режиме равен `0x100000`. Мы можем проверить это с помощью gdb. Давайте запустим отладчик и поставим точку останова на адресе `1f`, который равен `0x100021`. Если всё верно, то мы увидим `0x100021` в регистре `ebp`:

```
$ gdb
(gdb)$ target remote :1234
Remote debugging using :1234
0x0000fff0 in ?? ()
(gdb)$ br *0x100022
Breakpoint 1 at 0x100022
(gdb)$ c
Continuing.

Breakpoint 1, 0x00100022 in ?? ()
(gdb)$ i r
eax            0x18    0x18
ecx            0x0    0x0
edx            0x0    0x0
ebx            0x0    0x0
esp            0x144a8    0x144a8
ebp            0x100021    0x100021
esi            0x142c0    0x142c0
edi            0x0    0x0
eip            0x100022    0x100022
eflags         0x46    [ PF ZF ]
cs             0x10    0x10
ss             0x18    0x18
ds             0x18    0x18
es             0x18    0x18
fs             0x18    0x18
gs             0x18    0x18
```

Если мы выполним следующую инструкцию, `subl $1b, %ebp`, мы увидим следующее:

```
(gdb) nexti
...
...
...
ebp            0x100000    0x100000
...
...
...
```

Да, всё верно. Адрес `startup_32` равен `0x100000`. После того как мы узнали адрес метки `startup_32`, мы можем начать подготовку к переходу в long mode. Наша следующая цель - настроить стек и убедится в том, что CPU поддерживает long mode и SSE.

Настройка стека и проверка CPU

Мы не могли настроить стек, пока не знали адрес метки `startup_32`. Мы можем представить себе стек как массив, и регистр указателя стека `esp` должен указывать на конец этого массива. Конечно, мы можем определить массив в нашем коде, но мы должны знать его фактический адрес, чтобы правильно настроить указатель стека. Давайте посмотрим на код:

```
    movl    $boot_stack_end, %eax
    addl    %ebp, %eax
```

```
    movl    %eax, %esp
```

Метка `boot_stack_end` определена в arch/x86/boot/compressed/head_64.S и расположена в секции .bss:

```
    .bss
    .balign 4
boot_heap:
    .fill BOOT_HEAP_SIZE, 1, 0
boot_stack:
    .fill BOOT_STACK_SIZE, 1, 0
boot_stack_end:
```

Прежде всего, мы помещаем адрес `boot_stack_end` в регистр `eax`, т.е регистр `eax` содержит адрес `0x0 + boot_stack_end`. Чтобы получить реальный адрес `boot_stack_end`, нам нужно добавить реальный адрес `startup_32`. Как вы помните, мы нашли этот адрес выше и поместили его в регистр `ebp`. В итоге регистр `eax` будет содержать реальный адрес `boot_stack_end` и нам просто нужно поместить его в указатель стека.

После того как мы создали стек, следующим шагом является проверка CPU. Так как мы собираемся перейти в `long mode`, нам необходимо проверить, поддерживает ли CPU `long mode` и `SSE`. Мы будем делать это с помощью вызова функции `verify_cpu`:

```
    call    verify_cpu
    testl    %eax, %eax
    jnz    no_longmode
```

Она определена в arch/x86/kernel/verify_cpu.S и содержит пару вызовов инструкции CPUID. Данная инструкция используется для получения информации о процессоре. В нашем случае она проверяет поддержку `long mode` и `SSE` и с помощью регистра `eax` возвращает `0` в случае успеха или `1` в случае неудачи.

Если значение `eax` не равно нулю, то мы переходим на метку `no_longmode`, которая останавливает CPU вызовом инструкции `hlt` до тех пор, пока не произойдёт аппаратное прерывание:

```
no_longmode:
1:
    hlt
    jmp     1b
```

Если значение `eax` равно нулю, то всё в порядке и мы можем продолжить.

Расчёт адреса релокации

Следующим шагом является вычисление адреса релокации для декомпрессии, если это необходимо. Мы уже знаем, что базовый адрес 32-битной точки входа в ядро Linux - `0x100000`, но это 32-битная точка входа. Базовый адрес ядра по умолчанию определяется значением параметра конфигурации ядра `CONFIG_PHYSICAL_START`. Его значение по умолчанию `0x1000000` или `16 Мб`. Основная проблема заключается в том, что если происходит краш ядра, разработчик должен иметь `rescue ядро` ("спасательное" ядро) для kdump, которое сконфигурировано для загрузки из другого адреса. Для решения этой проблемы ядро Linux предоставляет специальный параметр конфигурации - `CONFIG_RELOCATABLE`. Как вы можете прочесть в документации ядра:

```
Это создает образ ядра, который сохраняет информацию о релокации
поэтому он может быть загружен где-либо, кроме стандартного 1 Мб.

Примечание: Если CONFIG_RELOCATABLE=y, то ядро запускается с адреса,
на который он был загружен, а физический адрес времени компиляции
(CONFIG_PHYSICAL_START) используется как минимальная локация.
```

Проще говоря, это означает, что ядро с той же конфигурацией может загружаться с разных адресов. С технической точки зрения это делается путём компиляции декомпрессора как адресно-независимого кода. Если мы посмотрим на arch/x86/boot/compressed/Makefile, то мы увидим, что декомпрессор действительно скомпилирован с флагом `-fPIC` :

```
KBUILD_CFLAGS += -fno-strict-aliasing -fPIC
```

Когда мы используем адресно-независимый код, адрес получается путём добавления адресного поля инструкции и значения счётчика команд программы. Код, использующий подобную адресацию, возможно загрузить с любого адреса. Вот почему мы должны были получить реальный физический адрес `startup_32` . Давайте вернёмся к коду ядра Linux. Наша текущая цель состоит в том, чтобы вычислить адрес, на который мы можем переместить ядро для декомпрессии. Расчёт этого адреса зависит от параметра конфигурации ядра `CONFIG_RELOCATABLE` . Давайте посмотрим на код:

```
#ifdef CONFIG_RELOCATABLE
    movl    %ebp, %ebx
    movl    BP_kernel_alignment(%esi), %eax
    decl    %eax
    addl    %eax, %ebx
    notl    %eax
    andl    %eax, %ebx
    cmpl    $LOAD_PHYSICAL_ADDR, %ebx
    jge    1f
#endif
    movl    $LOAD_PHYSICAL_ADDR, %ebx
```

Следует помнить, что регистр `ebp` содержит физический адрес метки `startup_32` . Если параметр `CONFIG_RELOCATABLE` включён во время конфигурации ядра, то мы помещаем этот адрес в регистр `ebx` , выравниваем по границе, кратной `2 Мб` и сравниваем его со значением `LOAD_PHYSICAL_ADDR` . `LOAD_PHYSICAL_ADDR` является макросом, определённым в arch/x86/include/asm/boot.h и выглядит следующим образом:

```
#define LOAD_PHYSICAL_ADDR ((CONFIG_PHYSICAL_START \
                + (CONFIG_PHYSICAL_ALIGN - 1)) \
                & ~(CONFIG_PHYSICAL_ALIGN - 1))
```

Как мы можем видеть, он просто расширяет адрес до значения выравнивания `CONFIG_PHYSICAL_ALIGN` и представляет собой физический адрес, по которому будет загружено ядро. После сравнения `LOAD_PHYSICAL_ADDR` и значения регистра `ebx` , мы добавляем смещение от `startup_32` , по которому будет происходить декомпрессия образа ядра. Если во время компиляции параметр `CONFIG_RELOCATABLE` не включён, мы просто помещаем адрес по умолчанию и добавляем к нему `z_extract_offset` .

После всех расчётов у нас в распоряжении `ebp` , содержащий адрес, по которому будет происходить загрузка, и `ebx` , содержащий адрес, по которому ядро будет перемещено после декомпрессии. Но это еще не конец. Сжатый образ ядра должен быть перемещён в конец буфера декомпрессии, чтобы упростить вычисления местоположения, по которому ядро будет расположено позже:

```
1:
    movl    BP_init_size(%esi), %eax
    subl    $_end, %eax
    addl    %eax, %ebx
```

мы помещаем значение из `boot_params.BP_init_size` (или значение заголовка настройки ядра из `hdr.init_size`) в регистр `eax` . `BP_init_size` содержит наибольшее значение между сжатым и распакованным vmlinux. Затем мы вычитаем адрес символа `_end` из этого значения и добавляем результат вычитания в регистр `ebx` , который хранит базовый адрес для декомпрессии ядра.

Подготовка перед входом в long mode

Теперь, когда у нас есть базовый адрес, на который мы будем перемещать сжатое ядро, нам необходимо сделать последний шаг, прежде чем мы сможем перейти в 64-битный режим. Во-первых, нам необходимо обновить глобальную таблицу дескрипторов с 64-битными сегментами, потому что перемещаемое ядро может быть запущено по любому адресу ниже 512 Гб:

```
    addl    %ebp, gdt+2(%ebp)
    lgdt    gdt(%ebp)
```

Здесь мы настраиваем базовый адрес `глобальной таблицы дескрипторов` на адрес, где мы фактически загружены, и загружаем таблицу с помощью инструкции `lgdt`.

Чтобы понять магию смещений `gdt`, нам нужно взглянуть на определение `глобальной таблицы дескрипторов`. Мы можем найти его определение в том же файле исходного кода:

```
    .data
gdt64:
    .word    gdt_end - gdt
    .long    0
    .word    0
    .quad   0
gdt:
    .word    gdt_end - gdt
    .long    gdt
    .word    0
    .quad    0x00cf9a000000ffff    /* __KERNEL32_CS *//
    .quad    0x00af9a000000ffff    /* __KERNEL_CS */
    .quad    0x00cf92000000ffff    /* __KERNEL_DS */
    .quad    0x0080890000000000    /* Дескриптор TS */
    .quad   0x0000000000000000    /* Продолжение TS */
gdt_end:
```

Мы видим, что она расположена в секции `.data` и содержит пять дескрипторов: `32-битный` дескриптор для сегмента кода ядра, `64-битный` сегмент ядра, сегмент данных ядра и два дескриптора задач.

Мы уже загрузили `глобальную таблицу дескрипторов` в предыдущей части, и теперь мы делаем почти то же самое здесь, но теперь дескрипторы с `CS.L = 1` и `CS.D = 0` для выполнения в 64-битном режиме. Как мы видим, определение `gdt` начинается с двух байт: `gdt_end - gdt`, который представляет последний байт `gdt` или лимит таблицы. Следующие 4 байта содержат базовый адрес `gdt`.

После того как `глобальная таблица дескрипторов` загружена с помощью инструкции `lgdt`, нам необходимо включить режим PAE, поместив значение регистра `cr4` в `eax`, установить в нём пятый бит и загрузить его снова в `cr4`:

```
    movl    %cr4, %eax
    orl    $X86_CR4_PAE, %eax
    movl    %eax, %cr4
```

Мы почти закончили все подготовки перед входом в 64-битный режим. Последний шаг заключается в создании таблицы страниц, но прежде чем сделать это, необходимо рассказать о long mode

Long mode

Long mode - нативный режим для процессоров x86_64. Прежде всего посмотрим на некоторые различия между `x86_64` и `x86`.

`64-битный` режим предоставляет следующие особенности:

- 8 новых регистров общего назначения с `r8` по `r15` + все регистры общего назначения теперь 64-битные;
- 64-битный указатель инструкции - `RIP` ;
- Новый режим работы - Long mode;
- 64-битные адреса и операнды;
- Относительная адресация RIP (мы увидим пример этого в следующих частях).

Long mode является расширением унаследованного защищённого режима. Он состоит из двух подрежимов:

- 64-битный режим;
- режим совместимости.

Для переключения в `64-битный` режим необходимо сделать следующее:

- Включить PAE;
- Создать таблицу страниц и загрузить адрес таблицы страниц верхнего уровня в регистр `cr3` ;
- Включить `EFER.LME` ;
- Включить страничную организацию памяти.

Мы уже включили `PAE` путём установки бита `PAE` в регистре управления `cr4` . Наша следующая цель - создать структуру для страничной организации. Мы увидим это в следующем параграфе.

Ранняя инициализация таблицы страниц

Итак, мы уже знаем, что прежде чем мы сможем перейти в `64-битный` режим, необходимо создать таблицу страниц. Давайте посмотри на создание ранних `4 гигабайтных` загрузочных таблиц страниц.

ПРИМЕЧАНИЕ: я не буду описывать теорию виртуальной памяти. Если вам необходимо больше информации по виртуальной памяти, см. ссылки в конце этой части.

Ядро Linux использует `4 уровневую` страничную организацию, и в целом мы создадим 6 таблиц страниц:

- Одну таблицу `PML4 (карта страниц 4 уровня, Page Map Level 4)` с одной записью;
- Одну таблицу `PDP (указатель директорий страниц, Page Directory Pointer)` с четырьмя записями;
- Четыре таблицы директорий страниц с `2048` записями.

Давайте посмотрим на реализацию. Прежде всего, мы очищаем буфер для таблиц страниц в памяти. Каждая таблица имеет размер в `4096` байт, поэтому нам необходимо очистить `24` Кб буфера:

```
	leal	pgtable(%ebx), %edi
	xorl	%eax, %eax
	movl	$(BOOT_INIT_PGT_SIZE/4), %ecx
	rep	stosl
```

Мы помещаем адрес `pgtable + ebx` (вы помните, что `ebx` содержит адрес, по которому ядро будет перемещено после декомпрессии) в регистр `edi` , очищаем регистр `eax` и устанавливаем регистр `ecx` в `6144` .

Инструкция `rep stosl` записывает значение `eax` в `edi` , увеличивает значение в регистре `edi` на `4` и уменьшает значение в регистре `ecx` на `1` . Эта операция будет повторятся до тех пор, пока значение регистра `ecx` больше нуля. Вот почему мы установили `ecx` в `6144` (или `BOOT_INIT_PGT_SIZE/4`).

Структура `pgtable` определена в конце файла arch/x86/boot/compressed/head_64.S:

```
	.section ".pgtable","a",@nobits
	.balign 4096
pgtable:
	.fill BOOT_PGT_SIZE, 1, 0
```

Как мы видим, она находится в секции `.pgtable` и его размер зависит от опции конфигурации ядра `CONFIG_X86_VERBOSE_BOOTUP`:

```
#  ifdef CONFIG_X86_VERBOSE_BOOTUP
#   define BOOT_PGT_SIZE    (19*4096)
#  else /* !CONFIG_X86_VERBOSE_BOOTUP */
#   define BOOT_PGT_SIZE    (17*4096)
#  endif
# else /* !CONFIG_RANDOMIZE_BASE */
#  define BOOT_PGT_SIZE        BOOT_INIT_PGT_SIZE
# endif
```

После того как мы получили буфер для `pgtable`, мы можем начать с создания таблицы страниц верхнего уровня - `PML4` - следующим образом:

```
    leal    pgtable + 0(%ebx), %edi
    leal    0x1007 (%edi), %eax
    movl    %eax, 0(%edi)
```

Здесь мы снова помещаем относительный адрес `pgtable` в `ebx` или, другими словами, относительный адрес `startup_32` в регистр `edi`. Далее мы помещаем этот адрес со смещением `0x1007` в регистр `eax`. Смещение `0x1007` равно `4096` байтам, которые представляют собой размер `PML4` плюс `7`. `7` здесь представляет флаги `PML4`. В нашем случае это флаги `PRESENT+RW+USER`. В конечном счёте мы просто записали адрес первого элемента `PDP` в `PML4`.

Следующий шаг - создание четырёх записей `директории страниц` в таблице `указателя директорий страниц` с теми же флагами `PRESENT+RW+USE`:

```
    leal    pgtable + 0x1000(%ebx), %edi
    leal    0x1007(%edi), %eax
    movl    $4, %ecx
1:  movl    %eax, 0(%edi)
    addl    $0x00001000, %eax
    addl    $8, %edi
    decl    %ecx
    jnz    1b
```

Мы помещаем базовый адрес указателя директорий страниц, который равен `4096` или, другими словами, смещение `0x1000` от таблицы `pgtable` в `edi`, и адрес первой записи указателя директорий страниц в регистр `eax`. Значение `4`, помещённое в регистр `ecx`, будет счётчиком в следующем цикле, в котором мы записываем адрес первой записи таблицы указателя директорий страниц в регистр `edi`. После этого `edi` будет содержать адрес первой записи указателя директорий страниц с флагами `0x7`. Далее мы просто вычисляем адрес следующих записей указателя директорий страниц, где каждая запись имеет размер `8` байт, и записываем их адреса в `eax`. Последний шаг в создании страничной организации памяти - создание `2048` записей с `2 мегабайтными` страницами:

```
    leal    pgtable + 0x2000(%ebx), %edi
    movl    $0x00000183, %eax
    movl    $2048, %ecx
1:  movl    %eax, 0(%edi)
    addl    $0x00200000, %eax
    addl    $8, %edi
    decl    %ecx
    jnz    1b
```

Здесь мы делаем почти тоже самое, как и в предыдущем примере; все записи с флагами `$0x00000183`: `PRESENT + WRITE + MBZ`. В итоге мы будем иметь `2048` `2 мегабайтных` страниц:

```
>>> 2048 * 0x00200000
4294967296
```

или `4 гигабайтную` таблицу страниц. Мы закончили создание нашей ранней структуры таблицы страниц, которая отображает `4` Гб на память и теперь мы можем поместить адрес таблицы страниц верхнего уровня - `PML4` - в регистр управления `cr3` :

```
    leal    pgtable(%ebx), %eax
    movl    %eax, %cr3
```

На этом всё. Все подготовки завершены и теперь мы можем перейти в long mode.

Переход в 64-битный режим

В первую очередь нам нужно установить флаг `EFER.LME` в MSR, равный `0xC0000080` :

```
    movl    $MSR_EFER, %ecx
    rdmsr
    btsl    $_EFER_LME, %eax
    wrmsr
```

Здесь мы помещаем флаг `MSR_EFER` (который определён в arch/x86/include/uapi/asm/msr-index.h) в регистр `ecx` и вызываем инструкцию `rdmsr` , которая считывает регистр MSR. После выполнения `rdmsr` , полученные данные будут находится в `edx:eax` , которые будут зависеть от значения `ecx` . Далее мы проверяем бит `EFER_LME` инструкцией `btsl` и с помощью инструкции `wrmsr` записываем данные из `eax` в регистр `MSR` .

На следующем шаге мы помещаем адрес сегмента кода ядра в стек (мы определили его в GDT) и помещаем адрес функции `startup_64` в `eax` .

```
    pushl    $__KERNEL_CS
    leal    startup_64(%ebp), %eax
```

После этого мы помещаем адрес в стек и включаем поддержку страничной организации путём установки битов `PG` и `PE` в регистре `cr0` :

```
    pushl    %eax
    movl    $(X86_CR0_PG | X86_CR0_PE), %eax
    movl    %eax, %cr0
```

и выполняем инструкцию:

```
lret
```

Вы должны помнить, что на предыдущем шаге мы поместили адрес функции `startup_64` в стек, и после инструкции `lret` , CPU извлекает адрес и переходит по нему.

После всего этого, мы, наконец, в 64-битном режиме:

```
    .code64
    .org 0x200
ENTRY(startup_64)
....
....
....
```

На этом всё!

Заключение

Это конец четвёртой части о процессе загрузки ядра Linux. В следующей части мы увидим декомпрессию ядра и многое другое.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- Защищённый режим
- Документация для разработчиков ПО на архитектуре Intel® 64 и IA-32
- GNU компоновщик
- SSE
- Страничная организация памяти (Википедия)
- Моделезависимый регистр
- Инструкция .fill
- Предыдущая часть
- Страничная организация памяти (OSDEV)
- Системы страничной организации памяти
- Пособие по страничной организации на x86

Процесс загрузки ядра. Часть 5.

Декомпрессия ядра

Это пятая часть серии `Процесса загрузки ядра`. Мы видели переход в 64-битный режим в предыдущей части и в этой части мы продолжим с этого момента. Прежде чем мы перейдём к коду ядра, мы увидим последние шаги: подготовку к декомпрессии ядра, перемещение и, непосредственно, декомпрессию ядра. Итак... давайте снова погрузимся в код ядра.

Подготовка к декомпрессии ядра

Мы остановились прямо перед переходом к `64-битной` точке входа - `startup_64`, расположенной в arch/x86/boot/compressed/head_64.S. В предыдущей части мы уже видели переход к `startup_64` в `startup_32`:

```
    pushl    $__KERNEL_CS
    leal    startup_64(%ebp), %eax
    ...
    ...
    ...
    pushl    %eax
    ...
    ...
    ...
    lret
```

Так как мы загрузили новую `глобальную таблицу дескрипторов`, и был переход CPU в другой режим (в нашем случае в `64-битный` режим), мы можем видеть настройку сегментов данных в начале `startup_64`:

```
    .code64
    .org 0x200
ENTRY(startup_64)
    xorl    %eax, %eax
    movl    %eax, %ds
    movl    %eax, %es
    movl    %eax, %ss
    movl    %eax, %fs
    movl    %eax, %gs
```

Все сегментные регистры, кроме регистра `cs`, теперь сброшены после того как мы перешли в `long mode`.

Следующий шаг - вычисление разницы между адресом, по которому скомпилировано ядро, и адресом, по которому оно было загружено:

```
#ifdef CONFIG_RELOCATABLE
    leaq    startup_32(%rip), %rbp
    movl    BP_kernel_alignment(%rsi), %eax
    decl    %eax
    addq    %rax, %rbp
    notq    %rax
    andq    %rax, %rbp
    cmpq    $LOAD_PHYSICAL_ADDR, %rbp
    jge    1f
#endif
    movq    $LOAD_PHYSICAL_ADDR, %rbp
1:
    leaq    z_extract_offset(%rbp), %rbx
```

`rbp` содержит начальный адрес распакованного ядра и после выполнения этого кода регистр `rbx` будет содержать адрес релокации ядра для декомпрессии. Такой код мы уже видели в `startup_32` (вы можете прочитать об этом в предыдущей части - Расчёт адреса релокации), но нам снова нужно вычислить его, поскольку загрузчик может использовать 64-битный протокол загрузки и в этом случае `startup_32` просто не будет выполнен.

На следующем шаге мы видим установку указателя стека, сброс регистра флагов и установку `GDT` заново из-за того, что в случае `64-битного` протокола `32-битный` сегмент кода может быть проигнорирован загрузчиком:

```
    leaq    boot_stack_end(%rbx), %rsp

    leaq    gdt(%rip), %rax
    movq    %rax, gdt64+2(%rip)
    lgdt    gdt64(%rip)

    pushq    $0
    popfq
```

Если вы посмотрите на исходный код ядра Linux после команды `lgdt gdt64(%rip)`, вы увидите, что есть некоторый дополнительный код. Этот код необходим для включения пятиуровневой страничной организации, в случае необходимости. В этой книге мы рассмотрим только четырёхуровневую страничную организацию, поэтому этот код будет проигнорирован.

Как вы можете видеть выше, регистр `rbx` содержит начальный адрес кода декомпрессора ядра, и мы помещаем этот адрес со смещением `boot_stack_end` в регистр `rsp`, который представляет указатель на вершину стека. После этого шага стек будет корректным. Вы можете найти определение `boot_stack_end` в конце arch/x86/boot/compressed/head_64.S:

```
    .bss
    .balign 4
boot_heap:
    .fill BOOT_HEAP_SIZE, 1, 0
boot_stack:
    .fill BOOT_STACK_SIZE, 1, 0
boot_stack_end:
```

Он расположен в конце секции `.bss`, прямо перед таблицей `.pgtable`. Если вы посмотрите сценарий компоновщика arch/x86/boot/compressed/vmlinux.lds.S, вы найдёте определения `.bss` и `.pgtable`.

После того как стек был настроен, мы можем скопировать сжатое ядро по адресу, который мы получили выше после вычисления адреса релокации распакованного ядра. Прежде чем перейти к деталям, давайте посмотрим на этот ассемблерный код:

```
    pushq    %rsi
    leaq    (_bss-8)(%rip), %rsi
    leaq    (_bss-8)(%rbx), %rdi
    movq    $_bss, %rcx
    shrq    $3, %rcx
    std
    rep    movsq
    cld
    popq    %rsi
```

Прежде всего, мы помещаем `rsi` в стек. Нам нужно сохранить значение `rsi`, потому что теперь этот регистр хранит указатель на `boot_params`, которая является структурой режима реальных адресов, содержащая связанные с загрузкой данные (вы должны помнить эту структуру, мы заполняли её в начале кода настройки ядра). В конце этого кода мы снова восстановим указатель на `boot_params` в `rsi`.

Следующие две инструкции `leaq` вычисляют эффективные адреса `rip` и `rbx` со смещением `_bss - 8` и помещают их в `rsi` и `rdi`. Зачем мы вычисляем эти адреса? На самом деле сжатый образ ядра находится между этим кодом копирования (от `startup_32` до текущего кода) и кодом декомпрессии. Вы можете проверить это, посмотрев сценарий компоновщика - arch/x86/boot/compressed/vmlinux.lds.S:

```
    . = 0;
    .head.text : {
        _head = . ;
        HEAD_TEXT
        _ehead = . ;
    }
    .rodata..compressed : {
        *(.rodata..compressed)
    }
    .text :    {
        _text = .;     /* Text */
        *(.text)
        *(.text.*)
        _etext = . ;
    }
```

Обратите внимание, что секция `.head.text` содержит `startup_32`. Вы можете помнить это из предыдущей части:

```
    __HEAD
    .code32
ENTRY(startup_32)
...
...
...
```

Секция `.text` содержит код декомпрессии:

```
    .text
relocated:
...
...
...
/*
 * Делает декомпрессию и переходит на новое ядро.
 */
...
```

`.rodata..compressed` содержит сжатый образ ядра. Таким образом, `rsi` будет содержать абсолютный адрес `_bss - 8`, а `rdi` будет содержать относительный адрес релокации `_bss - 8`. Когда мы сохраняем эти адреса в регистрах, мы помещаем адрес `_bss` в регистр `rcx`. Как вы можете видеть в скрипте компоновщика `vmlinux.lds.S`, он находится в конце всех секций с кодом настройки/ядра. Теперь мы можем начать копирование данных из `rsi` в `rdi` по `8` байт с помощью инструкции `movsq`.

Обратите внимание на инструкцию `std` перед копированием данных: она устанавливает флаг `DF`, означающий, что `rsi` и `rdi` будут уменьшаться. Другими словами, мы будем копировать байты задом наперёд. В конце мы очищаем флаг `DF` с помощью инструкции `cld` и восстанавливаем структуру `boot_params` в `rsi`.

После релокации мы имеем адрес секции `.text` и совершаем переход по нему:

```
    leaq    relocated(%rbx), %rax
    jmp    *%rax
```

Последняя подготовка перед декомпрессией ядра

В предыдущем абзаце мы видели, что секция `.text` начинается с метки `relocated`. Первое, что она делает - очищает секцию `bss`:

```
    xorl    %eax, %eax
    leaq    _bss(%rip), %rdi
    leaq    _ebss(%rip), %rcx
    subq    %rdi, %rcx
    shrq    $3, %rcx
    rep    stosq
```

Нам нужно инициализировать секцию `.bss`, потому что скоро мы перейдём к коду на C. Здесь мы просто очищаем `eax`, помещаем адрес `_bss` в `rdi` и `_ebss` в `rcx`, и заполняем его нулями с помощью инструкции `rep stosq`.

В конце мы видим вызов функции `extract_kernel`:

```
    pushq    %rsi
    movq    %rsi, %rdi
    leaq    boot_heap(%rip), %rsi
    leaq    input_data(%rip), %rdx
    movl    $z_input_len, %ecx
    movq    %rbp, %r8
    movq    $z_output_len, %r9
    call    extract_kernel
    popq    %rsi
```

Мы снова устанавливаем `rdi` в указатель на структуру `boot_params` и сохраняем его в стек. В то же время мы устанавливаем `rsi` для указания на область, которая должа использоваться для распаковки ядра. Последним шагом является подготовка параметров `extract_kernel` и вызов этой функции для распаковки ядра. Функция `extract_kernel` определена в arch/x86/boot/compressed/misc.c и принимает шесть аргументов:

- `rmode` - указатель на структуру boot_params, которая заполнена загрузчиком или во время ранней инициализации ядра;
- `heap` - указатель на `boot_heap`, представляющий собой начальный адрес ранней загрузочной кучи;
- `input_data` - указатель на начало сжатого ядра или, другими словами, указатель на `arch/x86/boot/compressed/vmlinux.bin.bz2`;
- `input_len` - размер сжатого ядра;
- `output` - начальный адрес будущего распакованного ядра;
- `output_len` - размер распакованного ядра;

Все аргументы буду передаваться через регистры согласно двоичному интерфейсу приложений System V (ABI). Мы закончили подготовку и переходим к декомпрессии ядра.

Декомпрессия ядра

Как мы видели в предыдущем абзаце, функция `extract_kernel` определена arch/x86/boot/compressed/misc.c и содержит шесть аргументов. Эта функция начинается с инициализации видео/консоли, которую мы уже видели в предыдущих частях. Нам нужно сделать это ещё раз, потому что мы не знаем, находились ли мы в режиме реальных адресов, использовался ли загрузчик, или загрузчик использовал `32` или `64-битный` протокол загрузки.

После первых шагов инициализации мы сохраняем указатели на начало и конец свободной памяти:

```
free_mem_ptr     = heap;
free_mem_end_ptr = heap + BOOT_HEAP_SIZE;
```

где `heap` является вторым параметром функции `extract_kernel`, который мы получили в arch/x86/boot/compressed/head_64.S:

```
leaq    boot_heap(%rip), %rsi
```

Как вы видели выше, `boot_heap` определён как:

```
boot_heap:
    .fill BOOT_HEAP_SIZE, 1, 0
```

где `BOOT_HEAP_SIZE` - это макрос, который раскрывается в `0x10000` (`0x400000` в случае `bzip2` ядра) и представляет собой размер кучи.

После инициализации указателей кучи, следующий шаг - вызов функции `choose_random_location` из arch/x86/boot/compressed/kaslr.c. Как можно догадаться из названия функции, она выбирает ячейку памяти, в которой будет разархивирован образ ядра. Может показаться странным, что нам нужно найти или даже `выбрать` место для декомпрессии сжатого образа ядра, но ядро Linux поддерживает технологию kASLR, которая позволяет загрузить распакованное ядро по случайному адресу из соображений безопасности.

Мы не будем рассматривать рандомизацию адреса загрузки ядра Linux в этой части, но сделаем это в следующей части.

Теперь мы вернёмся к misc.c. После получения адреса для образа ядра мы должны были совершить некоторые проверки и убедиться в том, что полученный случайный адрес правильно выровнен и является корректным:

```
if ((unsigned long)output & (MIN_KERNEL_ALIGN - 1))
    error("Destination physical address inappropriately aligned");

if (virt_addr & (MIN_KERNEL_ALIGN - 1))
    error("Destination virtual address inappropriately aligned");

if (heap > 0x3fffffffffffUL)
    error("Destination address too large");

if (virt_addr + max(output_len, kernel_total_size) > KERNEL_IMAGE_SIZE)
    error("Destination virtual address is beyond the kernel mapping area");

if ((unsigned long)output != LOAD_PHYSICAL_ADDR)
    error("Destination address does not match LOAD_PHYSICAL_ADDR");

if (virt_addr != LOAD_PHYSICAL_ADDR)
    error("Destination virtual address changed when not relocatable");
```

После этого мы увидим знакомое сообщение:

```
Decompressing Linux...
```

и вызываем функцию `__decompress` :

```
__decompress(input_data, input_len, NULL, NULL, output, output_len, NULL, error);
```

которая будет распаковывать ядро. Реализация функции `__decompress` зависит от того, какой алгоритм декомпрессии был выбран во время компиляции:

```
#ifdef CONFIG_KERNEL_GZIP
#include "../../../../lib/decompress_inflate.c"
#endif
```

```
#ifdef CONFIG_KERNEL_BZIP2
#include "../../../../lib/decompress_bunzip2.c"
#endif

#ifdef CONFIG_KERNEL_LZMA
#include "../../../../lib/decompress_unlzma.c"
#endif

#ifdef CONFIG_KERNEL_XZ
#include "../../../../lib/decompress_unxz.c"
#endif

#ifdef CONFIG_KERNEL_LZO
#include "../../../../lib/decompress_unlzo.c"
#endif

#ifdef CONFIG_KERNEL_LZ4
#include "../../../../lib/decompress_unlz4.c"
#endif
```

После того как ядро распаковано, остаются две последние функции - `parse_elf` и `handle_relocations`. Основное назначение этих функций - переместить распакованный образ ядра в правильное место памяти. Дело в том, что декомпрессор распаковывает на месте, и нам всё равно нужно переместить ядро на правильный адрес. Как мы уже знаем, образ ядра является исполняемым файлом ELF, поэтому главной целью функции `parse_elf` является перемещение загружаемых сегментов на правильный адрес. Мы можем видеть загружаемые сегменты в выводе программы `readelf`:

```
readelf -l vmlinux

Elf file type is EXEC (Executable file)
Entry point 0x1000000
There are 5 program headers, starting at offset 64

Program Headers:
  Type           Offset             VirtAddr           PhysAddr
                 FileSiz            MemSiz              Flags  Align
  LOAD           0x0000000000200000 0xffffffff81000000 0x0000000001000000
                 0x0000000000893000 0x0000000000893000  R E    200000
  LOAD           0x0000000000a93000 0xffffffff81893000 0x0000000001893000
                 0x000000000016d000 0x000000000016d000  RW     200000
  LOAD           0x0000000000c00000 0x0000000000000000 0x0000000001a00000
                 0x00000000000152d8 0x00000000000152d8  RW     200000
  LOAD           0x0000000000c16000 0xffffffff81a16000 0x0000000001a16000
                 0x0000000000138000 0x000000000029b000  RWE    200000
```

Цель функции `parse_elf` - загрузить эти сегменты по адресу `output`, который мы получили с помощью функции `choose_random_location`. Эта функция начинается с проверки сигнатуры ELF:

```
Elf64_Ehdr ehdr;
Elf64_Phdr *phdrs, *phdr;

memcpy(&ehdr, output, sizeof(ehdr));

if (ehdr.e_ident[EI_MAG0] != ELFMAG0 ||
    ehdr.e_ident[EI_MAG1] != ELFMAG1 ||
    ehdr.e_ident[EI_MAG2] != ELFMAG2 ||
    ehdr.e_ident[EI_MAG3] != ELFMAG3) {
        error("Kernel is not a valid ELF file");
        return;
}
```

и если файл некорректный, функция выводит сообщение об ошибке и останавливается. Если же `ELF` файл корректный, мы просматриваем все заголовки из указанного `ELF` файла и копируем все загружаемые сегменты с правильным адресом, выровненным по 2 мегабайтам, в выходной буфер:

```
    for (i = 0; i < ehdr.e_phnum; i++) {
        phdr = &phdrs[i];

        switch (phdr->p_type) {
        case PT_LOAD:
#ifdef CONFIG_X86_64
            if ((phdr->p_align % 0x200000) != 0)
                error("Alignment of LOAD segment isn't multiple of 2MB");
#endif
#ifdef CONFIG_RELOCATABLE
            dest = output;
            dest += (phdr->p_paddr - LOAD_PHYSICAL_ADDR);
#else
            dest = (void *)(phdr->p_paddr);
#endif
            memmove(dest, output + phdr->p_offset, phdr->p_filesz);
            break;
        default:
            break;
        }
    }
```

С этого момента все загружаемые сегменты находятся в правильном месте.

Следующим шагом после функции `parse_elf` является вызов функции `handle_relocations`. Реализация этой функции зависит от опции конфигурации ядра `CONFIG_X86_NEED_RELOCS`, и если она включена, то эта функция корректирует адреса в образе ядра и вызывается только в том случае, если во время конфигурации ядра была включена опция конфигурации `CONFIG_RANDOMIZE_BASE`. Реализация функции `handle_relocations` достаточно проста. Эта функция вычитает значение `LOAD_PHYSICAL_ADDR` из значения базового адреса загрузки ядра и, таким образом, мы получаем разницу между тем, где ядро было слинковано для загрузки и тем, где оно было фактически загружено. После этого мы можем выполнить релокацию ядра, поскольку мы знаем фактический адрес, по которому было загружено ядро, адрес по которому оно было слинковано для запуска и таблицу релокации, которая находится в конце образа ядра.

После перемещения ядра мы возвращаемся из `extract_kernel` обратно в arch/x86/boot/compressed/head_64.S.

Адрес ядра находится в регистре `rax` и мы совершаем переход по нему:

```
jmp    *%rax
```

На этом всё. Теперь мы в ядре!

Заключение

Это конец пятой части процесса загрузки ядра Linux. Мы больше не увидим статей о загрузке ядра (возможны обновления этой и предыдущих статей), но будет много статей о других внутренних компонентах ядра.

В следующей главе будут описаны более подробные сведения о процессе загрузки ядра Linux, например рандомизация адреса загрузки и т.д.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- Рандомизация размещения адресного пространства
- initrd
- long mode
- bzip2
- Инструкция RDdRand
- Счётчик временных меток
- Программируемый интервальный таймер
- Предыдущий пост

Процесс загрузки ядра. Часть 6.

Введение

Это шестая часть серии `Процесса загрузки ядра`. В предыдущей части мы увидели конец процесса загрузки ядра. Но мы пропустили некоторые важные дополнительные детали.

Как вы помните, точкой входа ядра Linux является функция `start_kernel` из файла main.c, которая начинает выполнение по адресу `LOAD_PHYSICAL_ADDR`. Этот адрес зависит от параметра конфигурации ядра `CONFIG_PHYSICAL_START`, который по умолчанию равен `0x1000000`:

```
config PHYSICAL_START
    hex "Physical address where the kernel is loaded" if (EXPERT || CRASH_DUMP)
    default "0x1000000"
    ---help---
      This gives the physical address where the kernel is loaded.
      ...
      ...
      ...
```

Это значение может быть изменено во время конфигурации ядра, но также может быть выбрано случайно. Для этого во время конфигурации ядра должна быть включена опция `CONFIG_RANDOMIZE_BASE`.

В этом случае будет рандомизирован физический адрес, по которому будет загружен и распакован образ ядра Linux. В этой части рассматривается случай, когда эта опция включена и адрес загрузки образа ядра будет рандомизирован из соображений безопасности.

Инициализация таблиц страниц

Перед тем как декомпрессор ядра начнёт поиск случайного адреса из диапазона, по которому ядро будет распаковано и загружено, таблицы страниц, отображённые "один в один" (identity mapped page tables), должны быть инициализированы. Если загрузчик использует 16-битный или 32-битный протокол загрузки, у нас уже есть таблицы страниц. Но в любом случае нам могут понадобиться новые страницы по требованию, если декомпрессор ядра выберет диапазон памяти за их пределами. Вот почему нам нужно создать новые таблицы таблиц, отображённые "один в один".

Да, создание таблиц является одним из первых шагов во время рандомизации адреса загрузки. Но прежде чем мы это рассмотрим, давайте попробуем вспомнить, откуда мы пришли к этому вопросу.

В предыдущей части, мы увидели переход в long mode и переход к точке входа декомпрессора ядра - функции `extract_kernel`. Рандомизация начинается с вызова данной функции:

```
void choose_random_location(unsigned long input,
                            unsigned long input_size,
                            unsigned long *output,
                            unsigned long output_size,
                            unsigned long *virt_addr)
{}
```

Как мы можем видеть, эта функция принимает следующие пять параметров:

- `input`;
- `input_size`;
- `output`;

* `output_isze` ;
* `virt_addr` .

Попытаемся понять что это за параметры. Первый параметр, `input` , поступает из параметров функции `extract_kernel` , расположенной в файле arch/x86/boot/compressed/misc.c:

```
asmlinkage __visible void *extract_kernel(void *rmode, memptr heap,
                                          unsigned char *input_data,
                                          unsigned long input_len,
                                          unsigned char *output,
                                          unsigned long output_len)
{
  ...
  ...
  ...
  choose_random_location((unsigned long)input_data, input_len,
                         (unsigned long *)&output,
                         max(output_len, kernel_total_size),
                         &virt_addr);
  ...
  ...
  ...
}
```

Этот параметр передаётся из кода ассемблера:

```
leaq    input_data(%rip), %rdx
```

в файле arch/x86/boot/compressed/head_64.S. `input_data` генерируется маленькой программой mkpiggy. Если вы компилировали ядро Linux своими руками, вы можете найти сгенерированный этой программой файл, расположенный в `linux/arch/x86/boot/compressed/piggy.S` . В моём случае этот файл выглядит так:

```
.section ".rodata..compressed","a",@progbits
.globl z_input_len
z_input_len = 6988196
.globl z_output_len
z_output_len = 29207032
.globl input_data, input_data_end
input_data:
.incbin "arch/x86/boot/compressed/vmlinux.bin.gz"
input_data_end:
```

Как вы можете видеть, он содержит четыре глобальных символа. Первые два, `z_input_len` и `z_output_len` , являются размерами сжатого и несжатого `vmlinux.bin.gz` . Третий - это наш `input_data` и он указывает на образ ядра Linux в бинарном формате (все отладочные символы, комментарии и информация о релокации удаляются). И последний, `input_data_end` , указывает на конец сжатого образа ядра.

Таким образом, наш первый параметр функции `choose_random_location` является указателем на сжатый образ ядра, встроенный в объектный файл `piggy.o` .

Второй параметр функции `choose_random_location` - `z_input_len` , который мы уже видели.

Третий и четвёртый параметры функции `choose_random_location` - это адрес, по которому размещено распакованное ядро и размер образа распакованного ядра. Адрес, по которому будет размещён образ ядра, получен из arch/x86/boot/compressed/head_64.S и это адрес `startup_32` , выровненный по границе 2 мегабайт. Размер распакованного ядра также получен из `piggy.S` , как и `z_output_len` .

Последним параметром функции `choose_random_location` является виртуальный адрес физического адреса загрузки ядра. По умолчанию он совпадает с физическим адресом загрузки по умолчанию:

};

```
unsigned long virt_addr = LOAD_PHYSICAL_ADDR;
```

который зависит от конфигурации ядра:

```
#define LOAD_PHYSICAL_ADDR ((CONFIG_PHYSICAL_START \
                + (CONFIG_PHYSICAL_ALIGN - 1)) \
                & ~(CONFIG_PHYSICAL_ALIGN - 1))
```

Теперь посмотрим на реализацию функции `choose_random_location`. Она начинается с проверки опции `nokaslr` из командной строки ядра:

```
if (cmdline_find_option_bool("nokaslr")) {
    warn("KASLR disabled: 'nokaslr' on cmdline.");
    return;
}
```

и если параметр установлен, `choose_random_location` завершает свою работу и адрес загрузки ядра не будет рандомизирован. Связанные параметры командной строки можно найти в документации ядра:

```
kaslr/nokaslr [X86]

Включение/выключение базового смещения ASLR ядра и модуля
(рандомизация размещения адресного пространства), если оно встроено в ядро.
Если выбран CONFIG_HIBERNATION, kASLR отключён по умолчанию.
Если kASLR включён, спящий режим будет выключен.
```

Предположим, что мы не передали `nokaslr` в командную строку ядра, а также включён параметр конфигурации ядра `CONFIG_RANDOMIZE_BASE`. В этом случае мы добавляем флаг `kASLR` к флагам загрузки ядра:

```
boot_params->hdr.loadflags |= KASLR_FLAG;
```

и следующим шагом является вызов функции:

```
initialize_identity_maps();
```

расположенной в файле arch/x86/boot/compressed/kaslr_64.c. Эта функция начинается с инициализации экземпляра структуры `x86_mapping_info`:

```
mapping_info.alloc_pgt_page = alloc_pgt_page;
mapping_info.context = &pgt_data;
mapping_info.page_flag = __PAGE_KERNEL_LARGE_EXEC | sev_me_mask;
mapping_info.kernpg_flag = _KERNPG_TABLE;
```

Определение структуры `x86_mapping_info` расположено в файле arch/x86/include/asm/init.h:

```
struct x86_mapping_info {
    void *(*alloc_pgt_page)(void *);
    void *context;
    unsigned long page_flag;
    unsigned long offset;
    bool direct_gbpages;
    unsigned long kernpg_flag;
};
```

Эта структура предоставляет информацию об отображениях памяти. Как вы помните из предыдущей части, мы уже настроили начальные страницы с 0 до `4G` . На данный момент нам может потребоваться доступ к памяти выше `4G` для загрузки ядра в случайном месте. Таким образом, функция `initialize_identity_maps` выполняет инициализацию области памяти для возможной новой таблицы страниц. Прежде всего, давайте взглянем на определение структуры `x86_mapping_info` .

`alloc_pgt_page` - это функция обратного вызова, которая будет вызываться для выделения пространства под запись в таблице страниц. Поле `context` является экземпляром структуры `alloc_pgt_data` , которая в нашем случае будет использоваться для отслеживания выделенных таблиц страниц. Поля `page_flag` и `kernpg_flag` являются флагами страниц. Первый представляет флаги для записей `PMD` или `PUD` . Второе поле `kernpg_flag` представляет флаги для страниц ядра, которые позже можно переопределить. Поле `direct_gbpages` представляет поддержку больших страниц, а последнее поле, `offset` представляет смещение между виртуальными адресами ядра и физическими адресами до уровня `PMD` .

`alloc_pgt_page` просто проверяет, есть ли место для новой страницы, и выделяет новую страницу:

```
entry = pages->pgt_buf + pages->pgt_buf_offset;
pages->pgt_buf_offset += PAGE_SIZE;
```

в буфере из структуры:

```
struct alloc_pgt_data {
    unsigned char *pgt_buf;
    unsigned long pgt_buf_size;
    unsigned long pgt_buf_offset;
};
```

и возвращает адрес новой страницы. Последняя цель функции `initialize_identity_maps` заключается в инициализации `pgdt_buf_size` и `pgt_buf_offset` . Поскольку мы только в фазе инициализации, функция `initialze_identity_maps` устанавливает `pgt_buf_offset` в ноль:

```
pgt_data.pgt_buf_offset = 0;
```

и `pgt_data.pgt_buf_size` будет установлен в `77824` или `69632` в зависимости от того, какой протокол загрузки использует загрузчик (64-битный или 32-битный). Тоже самое и для `pgt_data.pgt_buf` . Если загрузчик загрузил ядро в `startup_32` , `pgdt_data.pgdt_buf` укажет на на конец таблицы страниц, которая уже была инициализирована в arch/x86/boot/compressed/head_64.S:

```
pgt_data.pgt_buf = _pgtable + BOOT_INIT_PGT_SIZE;
```

где `_pgtable` указывает на начало этой таблицы страниц _pgtable. В случае, если загрузчик использовал 64-битный протокол загрузки и загрузил ядро в `startup_64` , ранние таблицы страниц должны быть созданы самим загрузчиком и `_pgtable` будет просто перезаписан:

```
pgt_data.pgt_buf = _pgtable
```

После инициализации буфера для новых таблиц страниц мы можем вернуться к функции `select_random_location` .

Избежание зарезервированных диапазонов памяти

После того как таблицы страниц, отображённые "один в один", инициализированы, мы можем начать выбор случайного местоположения, по которому мы поместим распакованный образ ядра. Но, как вы можете догадаться, мы не можем выбрать абсолютно любой адрес. Существует зарезервированные области памяти. Эти адреса занимают некоторые важные вещи, например, initrd, командная строка ядра и т.д. Функция

```
mem_avoid_init(input, input_size, *output);
```

поможет нам это сделать. Все небезопасные области памяти будут собраны в массив:

```
struct mem_vector {
    unsigned long long start;
    unsigned long long size;
};

static struct mem_vector mem_avoid[MEM_AVOID_MAX];
```

Где `MEM_AVOID_MAX` находится в перечислении `mem_avoid_index`, который представляет собой различные типы зарезервированных областей памяти:

```
enum mem_avoid_index {
    MEM_AVOID_ZO_RANGE = 0,
    MEM_AVOID_INITRD,
    MEM_AVOID_CMDLINE,
    MEM_AVOID_BOOTPARAMS,
    MEM_AVOID_MEMMAP_BEGIN,
    MEM_AVOID_MEMMAP_END = MEM_AVOID_MEMMAP_BEGIN + MAX_MEMMAP_REGIONS - 1,
    MEM_AVOID_MAX,
};
```

Оба расположены в файле arch/x86/boot/compressed/kaslr.c.

Давайте посмотрим на реализацию функции `mem_avoid_init`. Основная цель этой функции - хранить информацию о зарезервированных областях памяти, описанных в перечислении `mem_avoid_index` в массиве `mem_avoid`, и создавать новые страницы для таких областей в нашем новом буфере, отображённом "один в один". Многочисленные части для функции `mem_avoid_index` аналогичны, давайте посмотрим на одну из них:

```
mem_avoid[MEM_AVOID_ZO_RANGE].start = input;
mem_avoid[MEM_AVOID_ZO_RANGE].size = (output + init_size) - input;
add_identity_map(mem_avoid[MEM_AVOID_ZO_RANGE].start,
         mem_avoid[MEM_AVOID_ZO_RANGE].size);
```

В начале функция `mem_avoid_init` пытается избежать области памяти, которая используется для текущей декомпрессии ядра. Мы заполняем запись из массива `mem_avoid` с указанием начала и размера такой области и вызываем функцию `add_identity_map`, которая должна создать страницы, отображённые "один в один", для этого региона. Функция `add_identity_map` определена в файле arch/x86/boot/compressed/kaslr_64.c:

```
void add_identity_map(unsigned long start, unsigned long size)
{
    unsigned long end = start + size;

    start = round_down(start, PMD_SIZE);
    end = round_up(end, PMD_SIZE);
    if (start >= end)
        return;

    kernel_ident_mapping_init(&mapping_info, (pgd_t *)top_level_pgt,
                  start, end);
}
```

Как мы можем видеть, она выравнивает область памяти по границе 2 мегабайт и проверяет заданные начальные и конечные адреса.

В конце она вызывает функцию `kernel_ident_mapping_init` из файла arch/x86/mm/ident_map.c и передаёт экземпляр `mapping_info`, который мы инициализировали ранее, адрес таблицы страниц верхнего уровня и адреса области памяти, для которой необходимо создать новое отображение "один в один".

Функция `kernel_ident_mapping_init` устанавливает флаги по умолчанию для новых страниц, если они не были заданы:

```
if (!info->kernpg_flag)
    info->kernpg_flag = _KERNPG_TABLE;
```

и начинает создание 2 мегабайтных (из-за бита `PSE` в `mapping_info.page_flag`) страничных записей (`PGD -> P4D -> PUD -> PMD` в случае пятиуровневых таблиц страниц или `PGD -> PUD -> PMD` в случае четырёхуровневых таблиц страниц), относящихся к указанным адресам.

```
for (; addr < end; addr = next) {
    p4d_t *p4d;

    next = (addr & PGDIR_MASK) + PGDIR_SIZE;
    if (next > end)
        next = end;

    p4d = (p4d_t *)info->alloc_pgt_page(info->context);
    result = ident_p4d_init(info, p4d, addr, next);

    return result;
}
```

Прежде всего, мы находим следующую запись `глобального каталога страниц` для данного адреса, и если она больше, чем `end` данной области памяти, мы устанавливаем её в `end`. После этого мы выделяем новую страницу с нашим обратным вызовом `x86_mapping_info`, который мы уже рассмотрели выше, и вызываем функцию `ident_p4d_init`. Функция `ident_p4d_init` будет делать то же самое, но для низкоуровневых каталогов страниц (`p4d` -> `pud` -> `pmd`).

На этом всё.

Новые страницы, связанные с зарезервированными адресами, находятся в наших таблицах страниц. Это не конец функции `mem_avoid_init`, но другие части схожи. Они просто создают страницы для initrd, командной строки ядра и т.д.

Теперь мы можем вернуться к функции `choose_random_location`.

Рандомизация физического адреса

После сохранения зарезервированных областей памяти в массиве `mem_avoid` и создания для них страниц, отображённых "один в один", мы выбираем минимальный доступный адрес для произвольного выбора области памяти:

```
min_addr = min(*output, 512UL << 20);
```

Он должен быть меньше чем `512` мегабайт. Значение `512` мегабайт было выбрано для того, чтобы избежать неизвестных вещей в нижней части памяти.

Следующим шагом будет выбор случайных физических и виртуальных адресов для загрузки ядра. Сначала физические адреса:

```
random_addr = find_random_phys_addr(min_addr, output_size);
```

Функция `find_random_phys_addr` определена в том же файле:

```
static unsigned long find_random_phys_addr(unsigned long minimum,
                                           unsigned long image_size)
{
    minimum = ALIGN(minimum, CONFIG_PHYSICAL_ALIGN);

    if (process_efi_entries(minimum, image_size))
        return slots_fetch_random();

    process_e820_entries(minimum, image_size);
    return slots_fetch_random();
}
```

Основная задача `process_efi_entries` - найти все подходящие диапазоны памяти в доступной для загрузки ядра памяти. Если ядро скомпилировано и запущено на системе без поддержки EFI, поиск областей памяти продолжиться в регионах e820. Все найденные области памяти будут сохранены в массиве:

```
struct slot_area {
    unsigned long addr;
    int num;
};

#define MAX_SLOT_AREA 100

static struct slot_area slot_areas[MAX_SLOT_AREA];
```

Для декомпрессии ядро выберет случайный индекс из этого массива. Этот выбор будет выполнен функцией `slots_fetch_random`. Основная задача функции `slots_fetch_random` заключается в выборе случайного диапазона памяти из массива `slot_areas` с помощью функции `kaslr_get_random_long`:

```
slot = kaslr_get_random_long("Physical") % slot_max;
```

Функция `kaslr_get_random_long` определена в файле arch/x86/lib/kaslr.c и просто возвращает случайное число. Обратите внимание, что случайное число будет получено разными способами, зависящими от конфигурации ядра (выбор случайного числа, основываясь на счётчике времени, rdrand и т.д.).

Рандомизация виртуального адреса

После того как декомпрессором ядра была выбрана случайная область памяти, для неё будут созданы новые страницы, отображённые "один в один":

```
random_addr = find_random_phys_addr(min_addr, output_size);

if (*output != random_addr) {
        add_identity_map(random_addr, output_size);
        *output = random_addr;
}
```

После этого `output` будет хранить базовый адрес области памяти, где будет распаковано ядро. Но на данный момент, как вы помните, мы рандомизировали только физический адрес. В случае архитектуры x86_64 виртуальный адрес также должен быть рандомизирован:

```
if (IS_ENABLED(CONFIG_X86_64))
    random_addr = find_random_virt_addr(LOAD_PHYSICAL_ADDR, output_size);
```

```
*virt_addr = random_addr;
```

В архитектуре, отличной от `x86_64`, случайный виртуальный адрес будет совпадать со случайным физическим. Функция `find_random_virt_addr` вычисляет количество диапазонов виртуальной памяти, которые могут содержать образ ядра, и вызывает `kaslr_get_random_long`, которую мы уже видели ранее, когда пытались найти случайный `физический` адрес.

Теперь мы имеет как физические базовые случайные адреса (`*output`), так и виртуальные (`*virt_addr`) случайные адреса для декомпрессии ядра.

На этом всё.

Заключение

Это конец шестой и последней части процесса загрузки ядра Linux. Мы больше не увидим статей о загрузке ядра (возможны обновления этой и предыдущих статей), но будет много статей о других внутренних компонентах ядра.

Следующая глава посвящена инициализации ядра, и мы увидим первые шаги в коде инициализации ядра Linux.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- Рандомизация размещения адресного пространства
- Протокол загрузки ядра Linux
- Long mode
- initrd
- Перечисляемый тип
- Четырёхуровневые таблицы страниц
- Пятиуровневые таблицы страниц
- EFI
- e820
- Счётчик времени
- rdrand
- x86_64
- Предыдущая часть

Процесс инициализации ядра

Здесь вы увидите несколько статей, которые описывают полный цикл инициализации ядра с первого шага после того, как ядро распаковано и до запуска ядром первого процесса.

Примечание: данные статьи не будут описанием всех шагов инициализации ядра. Здесь будет описана только общая часть ядра, без обработки прерываний, ACPI и многих других частей. Все части, которые я пропустил, будут описаны в других главах.

- Первые шаги после декомпрессии ядра - описывает первые шаги в ядре.
- Начальная обработка прерываний и исключений - описывает инициализацию начальных прерываний и начального обработчика ошибки страницы.
- Последние приготовления перед точкой входа в ядро - описывает последние приготовления перед вызовом `start_kernel`.
- Точка входа в ядро - описывает первые шаги в общем коде ядра.
- Продолжение архитектурно-зависимой инициализации - описывает архитектурно-зависимую инициализацию.
- Архитектурно-зависимая инициализация, снова... - описывает продолжение процесса архитектурно-зависимой инициализации.
- Конец архитектурно-зависимой инициализации, почти... - описывает конец `setup_arch`.
- Инициализация планировщика - описывает подготовку перед инициализацией и саму инициализацию планировщика.
- Инициализация RCU - описывает инициализацию RCU.
- Конец инициализации - последняя часть об инициализации ядра Linux.

Инициализация ядра. Часть 1.

Первые шаги в коде ядра

Предыдущая статья была последней частью главы процесса загрузки ядра Linux и теперь мы начинаем погружение в процесс инициализации. После того как образ ядра Linux распакован и помещён в нужное место, ядро начинает свою работу. Все предыдущие части описывают работу кода настройки ядра, который выполняет подготовку до того, как будут выполнены первые байты кода ядра Linux. Теперь мы находимся в ядре, и все части этой главы будут посвящены процессу инициализации ядра, прежде чем оно запустит процесс с помощью pid `1`. Есть ещё много вещей, который необходимо сделать, прежде чем ядро запустит первый `init` процесс. Мы начнём с точки входа в ядро, которая находится в arch/x86/kernel/head_64.S и будем двигаться дальше и дальше. Мы увидим первые приготовления, такие как инициализацию начальных таблиц страниц, переход на новый дескриптор в пространстве ядра и многое другое, прежде чем увидим запуск функции `start_kernel` в init/main.c.

В последней части предыдущей главы мы остановились на инструкции jmp из ассемблерного файла arch/x86/boot/compressed/head_64.S:

```
jmp    *%rax
```

В данный момент регистр `rax` содержит адрес точки входа в ядро Linux, который был получен в результате вызова функции `decompress_kernel` из файла arch/x86/boot/compressed/misc.c. Итак, наша последняя инструкция в коде настройки ядра - это переход на точку входа. Мы уже знаем, где определена точка входа ядра Linux, поэтому мы можем начать изучать, что делает ядро Linux после запуска.

Первые шаги в ядре

Хорошо, мы получили адрес распакованного образа ядра с помощью функции `decompress_kernel` в регистр `rax`. Как мы уже знаем, начальная точка распакованного образа ядра находится в файле arch/x86/kernel/head_64.S, а также в его начале можно увидеть следующие определения:

```
    .text
    __HEAD
    .code64
    .globl startup_64
startup_64:
    ...
    ...
    ...
```

Мы можем видеть определение подпрограммы `startup_64` в секции `__HEAD`, которая является просто макросом, раскрывающимся до определения исполняемой секции `.head.text`:

```
#define __HEAD        .section    ".head.text","ax"
```

Определение данной секции расположено в скрипте компоновщика arch/x86/kernel/vmlinux.lds.S:

```
.text : AT(ADDR(.text) - LOAD_OFFSET) {
    _text = .;
    ...
    ...
    ...
```

```
} :text = 0x9090
```

Помимо определения секции `.text` из скрипта компоновщика, мы можем понять виртуальные и физические адреса по умолчанию. Обратите внимание, что адрес `_text` - это счётчик местоположения, определённый как:

```
. = __START_KERNEL;
```

для x86_64. Определение макроса `__START_KERNEL` находится в заголовочном файле arch/x86/include/asm/page_types.h и представлен суммой базового виртуального адреса отображения ядра и физического начала:

```
#define __START_KERNEL    (__START_KERNEL_map + __PHYSICAL_START)

#define __PHYSICAL_START  ALIGN(CONFIG_PHYSICAL_START, CONFIG_PHYSICAL_ALIGN)
```

Или другими словами:

- Базовый физический адрес ядра Linux - `0x1000000` ;
- Базовый виртуальный адрес ядра Linux - `0xffffffff81000000` .

После того как мы очистили конфигурацию CPU, мы вызываем функцию `__startup_64` , которая определена в [arch/x86/kernel/head64.c](https://github.com/torvalds/linux/blob/master/arch/x86/kernel/head64.c:

```
    leaq    _text(%rip), %rdi
    pushq    %rsi
    call    __startup_64
    popq    %rsi
```

```
unsigned log __head __startup_64(unsigned long physaddr,
                 struct boot_params *bp)
{
    unsigned long load_delta, *p;
    unsigned long pgtable_flags;
    pgdval_t *pgd;
    p4dval_t *p4d;
    pudval_t *pud;
    pmdval_t *pmd, pmd_entry;
    pteval_t *mask_ptr;
    bool la57;
    int i;
    unsigned int *next_pgt_ptr;
    ...
    ...
    ...
}
```

Поскольку kASLR включен, адрес `start_64` может отличаться от адреса, скомпилированного для запуска, поэтому нам нужно вычислить дельту с помощью следующего кода:

```
    load_delta = physaddr - (unsigned long)(_text - __START_KERNEL_map);
```

В результате `load_delta` содержит дельту между адресом, скомпилированным для запуска, и текущим адресом.

После того как мы получили дельту, мы проверяем правильность выравнивания адреса `_text` по `2` мегабайтам. Мы сделаем это с помощью следующего кода:

```
    if (load_delta & ~PMD_PAGE_MASK)
        for (;;);
```

Если адрес `_text` не выровнен по `2` мегабайтам, мы входим в бесконечный цикл. `PMD_PAGE_MASK` указывает маску для `промежуточного каталога страниц` (см. страничную организацию памяти) и определён как:

```
#define PMD_PAGE_MASK           (~(PMD_PAGE_SIZE-1))
```

где макрос `PMD_PAGE_SIZE` определён как:

```
#define PMD_PAGE_SIZE           (_AC(1, UL) << PMD_SHIFT)
#define PMD_SHIFT       21
```

Размер `PMD_PAGE_SIZE` можно легко вычислить - он составляет `2` мегабайта.

Если поддержка SME#Enhanced_security_and_virtualization_support) включена, мы активируем её и включаем маску шифрования SME в `load_delta` :

```
    sme_enable(bp);
    load_delta += sme_get_me_mask();
```

Хорошо, мы сделали некоторые начальные проверки, и теперь можем двигаться дальше.

Исправление базовых адресов таблиц страниц

На следующем этапе мы исправляем физические адреса в таблице страниц:

```
    pgd = fixup_pointer(&early_top_pgt, physaddr);
    pud = fixup_pointer(&level3_kernel_pgt, physaddr);
    pmd = fixup_pointer(level2_fixmap_pgt, physaddr);
```

Давайте рассмотрим определение функции `fixup_pointer` , которая возвращает физический адрес переданного аргумента:

```
static void __head *fixup_pointer(void *ptr, unsigned long physaddr)
{
    return ptr - (void *)_text + (void *)physaddr;
}
```

Затем мы сосредоточимся на `early_top_pgt` и других табличных символах, которые мы видели выше. Давайте попробуем понять, что означают эти символы. Прежде всего посмотрим на их определение:

```
NEXT_PAGE(early_top_pgt)
    .fill    512,8,0
    .fill    PTI_USER_PGD_FILL,8,0

NEXT_PAGE(level3_kernel_pgt)
    .fill    L3_START_KERNEL,8,0
    .quad    level2_kernel_pgt - __START_KERNEL_map + _KERNPG_TABLE_NOENC
    .quad    level2_fixmap_pgt - __START_KERNEL_map + _PAGE_TABLE_NOENC

NEXT_PAGE(level2_kernel_pgt)
    PMDS(0, __PAGE_KERNEL_LARGE_EXEC,
        KERNEL_IMAGE_SIZE/PMD_SIZE)

NEXT_PAGE(level2_fixmap_pgt)
    .fill    506,8,0
    .quad    level1_fixmap_pgt - __START_KERNEL_map + _PAGE_TABLE_NOENC
    .fill    5,8,0
```

```
NEXT_PAGE(level1_fixmap_pgt)
    .fill    512,8,0
```

Выглядит сложно, но на самом деле это не так. Прежде всего, давайте посмотрим на `early_top_pgt`. Он начинается с `4096` нулевых байтов (или `8192` байт если включён `CONFIG_PAGE_TABLE_ISOLATION`), это означает, что мы не используем первые `511` записей. После этого мы видим одну запись `level3_kernel_pgt`. В начале его определения мы видим, что он заполнен `4080` байтами нулей (`L3_START_KERNEL` равен `510`). Впоследствии он хранит две записи, которые отображают пространство ядра. Обратите внимание, что мы вычитаем `__START_KERNEL_map` из `level2_kernel_pgt` и `level2_fixmap_pgt`. Как известно, `__START_KERNEL_map` является базовым виртуальным адресом текстового сегмента ядра, поэтому, если мы вычтем `__START_KERNEL_map`, мы получим физические адреса `level2_kernel_pgt` и `level2_fixmap_pgt`.

```
#define _KERNPG_TABLE_NOENC   (_PAGE_PRESENT | _PAGE_RW | _PAGE_ACCESSED | \
                   _PAGE_DIRTY)
#define _PAGE_TABLE_NOENC     (_PAGE_PRESENT | _PAGE_RW | _PAGE_USER | \
                   _PAGE_ACCESSED | _PAGE_DIRTY)
```

`level2_kernel_pgt` - это запись в таблице страниц, содержащая указатель на промежуточный каталог страниц, которая отображает пространство ядра. Она вызывает макрос `PDMS`, который создает `512` мегабайт из `__START_KERNEL_map` для `.text` ядра (после того как эти `512` мегабайт будут областью памяти модуля).

`level2_fixmap_pgt` - это виртуальные адреса, которые могут ссылаться на любые физические адреса даже в пространстве ядра. Они представлены `4048` байтами нулей, значением `level1_fixmap_pgt`, `8` мегабайтами, зарезервированными для отображения `vsyscalls` и `2` мегабайта пустого пространства.

Вы можете больше узнать об этом в статье страничная организация памяти.

Теперь, после того как мы увидели определения этих символов, вернёмся к коду. Мы инициализируем последнюю запись `pgd` с помощью `level3_kernel_pgt`:

```
pgd[pgd_index(__START_KERNEL_map)] = level3_kernel_pgt - __START_KERNEL_map + _PAGE_TABLE_NOENC;
```

Все адреса `p*d` могут быть неверными, если `startup_64` не равен адресу по умолчанию - `0x1000000`. Вы должны помнить, что `load_delta` содержит дельта между адресом метки `startup_64`, который был получен во время компоновки ядра и фактическим адресом. Таким образом, мы добавляем дельту к некоторым записям `p*d`:

```
    pgd[pgd_index(__START_KERNEL_map)] += load_delta;
    pud[510] += load_delta;
    pud[511] += load_delta;
    pmd[506] += load_delta;
```

После этого у нас будет:

```
early_top_pgt[511] -> level3_kernel_pgt[0]
level3_kernel_pgt[510] -> level2_kernel_pgt[0]
level3_kernel_pgt[511] -> level2_fixmap_pgt[0]
level2_kernel_pgt[0]   -> 512 Мб, отображённые на ядро
level2_fixmap_pgt[506] -> level1_fixmap_pgt
```

Обратите внимание, что мы не исправили базовый адрес `early_top_pgt` и некоторых других каталогов таблицы страниц, потому что мы увидим это во время построения/заполнения структур этих таблиц страниц. После исправления базовых адресов таблиц страниц, мы можем приступить к их построению.

Настройка отображения "один в один" (identity mapping)

Теперь мы можем увидеть настройку отображения "один в один" начальных таблиц страниц. В страничной организации с отображением "один в один", виртуальные адреса идентичны физическими адресами. Давайте рассмотрим это подробнее. Прежде всего, мы заменим `pud` и `pmd` указателем на первую и вторую запись `early_dynamic_pgts` :

```
    next_pgt_ptr = fixup_pointer(&next_early_pgt, physaddr);
    pud = fixup_pointer(early_dynamic_pgts[(*next_pgt_ptr)++], physaddr);
    pmd = fixup_pointer(early_dynamic_pgts[(*next_pgt_ptr)++], physaddr);
```

Давайте посмотри на определение `early_dynamic_pgts` :

```
NEXT_PAGE(early_dynamic_pgts)
    .fill    512*EARLY_DYNAMIC_PAGE_TABLES,8,0
```

которая будет хранить временные таблицы страниц раннего ядра.

Затем мы инициализируем `pgtable_flags` , который позже будет использоваться при инициализации записей `p*d` :

```
    pgtable_flags = _KERNPG_TABLE_NOENC + sme_get_me_mask();
```

Функция `sme_get_me_mask` возвращает `sme_me_mask` , который был инициализирован в функции `sme_enable` .

Далее мы заполняем две записи `pgd` с помощью `pud` плюс `pgtable_flags` , который мы инициализировали ранее:

```
    i = (physaddr >> PGDIR_SHIFT) % PTRS_PER_PGD;
    pgd[i + 0] = (pgdval_t)pud + pgtable_flags;
    pgd[i + 1] = (pgdval_t)pud + pgtable_flags;
```

`PGDIR_SHFT` обозначате маску для бит глобального каталога страниц в виртуальном адресе. Здесь мы вычисляем по модулю `PTRS_PER_PGD` (который раскрывается до `512`), чтобы не получить доступ к индексу, превышающему `512` . Для всех типов каталогов страниц есть свой макрос:

```
#define PGDIR_SHIFT     39
#define PTRS_PER_PGD    512
#define PUD_SHIFT       30
#define PTRS_PER_PUD    512
#define PMD_SHIFT       21
#define PTRS_PER_PMD    512
```

Мы делаем почти то же самое:

```
    i = (physaddr >> PUD_SHIFT) % PTRS_PER_PUD;
    pud[i + 0] = (pudval_t)pmd + pgtable_flags;
    pud[i + 1] = (pudval_t)pmd + pgtable_flags;
```

Затем мы инициализируем `pmd_entry` и отфильтровываем неподдерживаемые биты `__PAGE_KERNEL_*` :

```
    pmd_entry = __PAGE_KERNEL_LARGE_EXEC & ~_PAGE_GLOBAL;
    mask_ptr = fixup_pointer(&__supported_pte_mask, physaddr);
    pmd_entry &= *mask_ptr;
    pmd_entry += sme_get_me_mask();
    pmd_entry += physaddr;
```

Далее мы заполняем все записи `pmd` , чтобы покрыть полный размер ядра:

```
    for (i = 0; i < DIV_ROUND_UP(_end - _text, PMD_SIZE); i++) {
        int idx = i + (physaddr >> PMD_SHIFT) % PTRS_PER_PMD;
```

```
        pmd[idx] = pmd_entry + i * PMD_SIZE;
    }
```

Затем мы исправляем виртуальные адреса текста+данных ядра. Обратите внимание, что мы можем записать недопустимые `pmd`, если ядро было перемещено (функция `cleanup_highmap` исправляет это вместе с отображениями вне `_end`).

```
    pmd = fixup_pointer(level2_kernel_pgt, physaddr);
    for (i = 0; i < PTRS_PER_PMD; i++) {
        if (pmd[i] & _PAGE_PRESENT)
            pmd[i] += load_delta;
    }
```

Далее мы удаляем маску шифрования памяти для получения истинного физического адреса (помните, что `load_delta` включает в себя маску):

```
    *fixup_long(&phys_base, physaddr) += load_delta - sme_get_me_mask();
```

`phys_base` должен соответствовать первой записи в `level2_kernel_pgt`.

В качестве последнего шага функции `__startup_64` мы зашифровываем ядро (если активен SME) и возвращаем маску шифрования SME, которая будет использоваться в качестве модификатора для начальной записи каталога страницы, запрограммированной в регистр `cr3`:

```
    sme_encrypt_kernel(bp);
    return sme_get_me_mask();
```

Теперь вернемся к ассемблерному коду. Мы готовимся к следующему разделу со следующим кодом:

```
    addq    $(early_top_pgt - __START_KERNEL_map), %rax
    jmp 1f
```

который добавляет физический адрес `early_top_pgt` к регистру `rax` и теперь регистр `rax` содержит сумму адреса и маски шифрования SME.

На данный момент это всё. Наша ранняя страничная структура настроена и нам нужно совершить последнее приготовление, прежде чем мы перейдем к точке входа в ядро.

Последнее приготовление перед переходом на точку входа в ядро

После перехода на метку `1` мы включаем `PAE`, `PGE` (Paging Global Extension) и помещаем содержимое `phys_base` (см. выше) в регистр `rax` и заполняем регистр `cr3`:

```
1:
    movl    $(X86_CR4_PAE | X86_CR4_PGE), %ecx
    movq    %rcx, %cr4

    addq    phys_base(%rip), %rax
    movq    %rax, %cr3
```

На следующем шаге мы проверяем, поддерживает ли процессор бит NX:

```
    movl    $0x80000001, %eax
```

```
    cpuid
    movl    %edx,%edi
```

Мы помещаем значение `0x80000001` в `eax` и выполняем инструкцию `cpuid` для получения расширенной информации о процессоре и битах. Полученный результат находится в регистре `edx`, который мы помещаем в `edi`.

Теперь мы помещаем `0xc0000080` (`MSR_EFER`) в `ecx` и вызываем инструкцию `rdmsr` для чтения моделезависимого регистра.

```
    movl    $MSR_EFER, %ecx
    rdmsr
```

Результат находится в `edx:eax`. Общий вид `EFER` следующий:

```
63                                                                  32
 ┌───────────────────────────────────────────────────────────────────┐
 |                                                                   |
 |                          Зарезервированный MBZ                    |
 |                                                                   |
 └───────────────────────────────────────────────────────────────────┘
31                        16  15      14      13   12  11   10  9  8 7  1   0
 ┌──────────────────────────┬───┬──────┬──────┬────┬───┬───┬───┬───┬───┬───┬───┐
 |                          | T |      |      |    |   |   |   |   |   |   |   |
 | Зарезервированный MBZ    | C | FFXSR| LMSLE|SVME|NXE|LMA|MBZ|LME|RAZ|SCE|
 |                          | E |      |      |    |   |   |   |   |   |   |   |
 └──────────────────────────┴───┴──────┴──────┴────┴───┴───┴───┴───┴───┴───┴───┘
```

Здесь мы не увидим все поля, но узнаем об этих и других `MSR` в специальной части. Когда мы считываем `EFER` в `edx:eax`, мы проверяем `_EFER_SCE` или нулевой бит, являющийся `System Call Extensions` с инструкцией `btsl` и устанавливаем его в единицу. С помощью бита `SCE` мы включаем инструкции `SYSCALL` и `SYSRET`. На следующем шаге мы проверяем 20 бит в регистре `edi`, который хранит результат `cpuid` (см. выше). Если `20` бит установлен (бит `NX`), мы просто записываем `EFER_SCE` в моделезависимый регистр.

```
    btsl    $_EFER_SCE, %eax
    btl    $20,%edi
    jnc     1f
    btsl    $_EFER_NX, %eax
    btsq    $_PAGE_BIT_NX,early_pmd_flags(%rip)
1:    wrmsr
```

Если бит NX поддерживается, мы включаем `_EFER_NX` и записываем в него с помощью инструкции `wrmsr`. После того как бит NX установлен, мы устанавливаем некоторые биты в регистре управления `cr0`:

```
    movl    $CR0_STATE, %eax
    movq    %rax, %cr0
```

в частности следующие биты:

- `X86_CR0_PE` - система в защищённом режиме;
- `X86_CR0_MP` - контролирует взаимодействие инструкций WAIT/FWAIT с помощью флага TS в CR0;
- `X86_CR0_ET` - на 386 позволяло указать, был ли внешний математический сопроцессор 80287 или 80387;
- `X86_CR0_NE` - позволяет включить внутреннюю x87 отчётность об ошибках с плавающей запятой, иначе включает PC-стиль x87 обнаружение ошибок;
- `X86_CR0_WP` - если установлен, CPU не может писать в страницы только для чтения, когда уровень привилегий равен 0;
- `X86_CR0_AM` - проверка выравнивания включена, если установлен AM и флаг AC (в регистре EFLAGS), а уровень

привелигий равен 3;
- `X86_CR0_PG` - включает страничную организацию.

Мы уже знаем, что для запуска любого кода и даже большего количества C кода из ассемблера, нам необходимо настроить стек. Как всегда, мы делаем это путём установки указателя стека на корректное место в памяти и сброса регистра флагов:

```
movq initial_stack(%rip), %rsp
pushq $0
popfq
```

Самое интересное здесь - `initial_stack` . Этот символ определён в файле arch/x86/kernel/head_64.S и выглядит следующим образом:

```
GLOBAL(initial_stack)
    .quad  init_thread_union + THREAD_SIZE - SIZEOF_PTREGS
```

Макрос `THREAD_SIZE` определён в arch/x86/include/asm/page_64_types.h и зависит от значения макроса `KASAN_STACK_ORDER` :

```
#ifdef CONFIG_KASAN
#define KASAN_STACK_ORDER 1
#else
#define KASAN_STACK_ORDER 0
#endif

#define THREAD_SIZE_ORDER       (2 + KASAN_STACK_ORDER)
#define THREAD_SIZE  (PAGE_SIZE << THREAD_SIZE_ORDER)
```

когда kasan отключён, а `PAGE_SIZE` равен `4096` байтам. Таким образом, `THREAD_SIZE` будет раскрыт до `16` килобайт и представляет собой размер стека потока. Почему `потока` ? Возможно, вы уже знаете, что каждый процесс может иметь родительский процесс и дочерний процессы. На самом деле родительский и дочерний процесс различаются в стеке. Для нового процесса выделяется новый стек ядра. В ядре Linux этот стек представлен объединением (union) со структурой `thread_info` .

`init_thread_union` представлен `thread_union` и определён в файле include/linux/sched.h:

```
union thread_union {
#ifndef CONFIG_ARCH_TASK_STRUCT_ON_STACK
    struct task_struct task;
#endif
#ifndef CONFIG_THREAD_INFO_IN_TASK
    struct thread_info thread_info;
#endif
    unsigned long stack[THREAD_SIZE/sizeof(long)];
};
```

где `CONFIG_THREAD_INFO_IN_TASK` - параметр конфигурации ядра, включённый для архитектуры `ia64` , а `CONFIG_THREAD_INFO_IN_TASK` - параметр конфигурации ядра, включённый для архитектуры `x86_64` . Таким образом, структура `thread_info` будет помещена в структуру `task_struct` вместо объединения `thread_union` .

`init_thread_union` расположен в файле include/asm-generic/vmlinux.lds.h как часть макроса `INIT_TASK_DATA` :

```
#define INIT_TASK_DATA(align)  \
    ...                    \
    init_thread_union = .; \
    ...
```

Данный макрос используется в arch/x86/kernel/vmlinux.lds.S следующим образом:

```
.data : AT(ADDR(.data) - LOAD_OFFSET) {
    ...
    INIT_TASK_DATA(THREAD_SIZE)
    ...
} :data
```

Теперь мы можем понять это выражение:

```
GLOBAL(initial_stack)
    .quad  init_thread_union + THREAD_SIZE - SIZEOF_PTREGS
```

где символ `initial_stack` указывает на начало массива `thread_union.stack` + `THREAD_SIZE` , который равен 16 килобайтам и - `SIZEOF_PTREGS` , который является соглашением, помогающее unwinder'у ядра надёжно обнаруживать конец стека.

После настройки начального загрузочного стека, необходимо обновить глобальную таблицу дескрипторов с помощью инструкции `lgdt` :

```
lgdt    early_gdt_descr(%rip)
```

где `early_gdt_descr` определён как:

```
early_gdt_descr:
    .word    GDT_ENTRIES*8-1
early_gdt_descr_base:
    .quad    INIT_PER_CPU_VAR(gdt_page)
```

Это необходимо, поскольку теперь ядро работает в нижних адресах пользовательского пространства, но вскоре ядро будет работать в своём собственном пространстве.

Теперь давайте посмотрим на определение `early_gdt_descr` . Макрос `GDT_ENTRIES` раскрывается до `32` , поэтому глобальная таблица дескрипторов содержит `32` записи для кода ядра, данных, сегментов локального хранилища потоков и т.д.

Теперь давайте посмотрим на определение `early_gdt_descr_base` . Структура `gdt_page` определена в arch/x86/include/asm/desc.h:

```
struct gdt_page {
    struct desc_struct gdt[GDT_ENTRIES];
} __attribute__((aligned(PAGE_SIZE)));
```

Она содержит одно поле `gdt` , которое является массивом структур `desc_struct` :

```
struct desc_struct {
        union {
                struct {
                        unsigned int a;
                        unsigned int b;
                };
                struct {
                        u16 limit0;
                        u16 base0;
                        unsigned base1: 8, type: 4, s: 1, dpl: 2, p: 1;
                        unsigned limit: 4, avl: 1, l: 1, d: 1, g: 1, base2: 8;
                };
        };
```

```
} __attribute__((packed));
```

который выглядит знакомым дескриптором `GDT` . Можно отметить, что структура `gdt_page` выровнена по `PAGE_SIZE` , равному `4096` байтам. Это значит, что `gdt` займёт одну страницу.

Теперь попробуем понять, что такое `INIT_PER_CPU_VAR` . `INIT_PER_CPU_VAR` это макрос, определённый в arch/x86/include/asm/percpu.h, который просто совершает конкатенацию `init_per_cpu__` с заданным параметром:

```
#define INIT_PER_CPU_VAR(var) init_per_cpu__##var
```

После того, как макрос `INIT_PER_CPU_VAR` будет раскрыт, мы будем иметь `init_per_cpu__gdt_page` . Мы можем видеть инициализацию `init_per_cpu__gdt_page` в скрипте компоновщика:

```
#define INIT_PER_CPU(x) init_per_cpu__##x = x + __per_cpu_load
INIT_PER_CPU(gdt_page);
```

После того как макросы `INIT_PER_CPU_VAR` и `INIT_PER_CPU` будут раскрыты до `init_per_cpu__gdt_page` мы получим смещение от `__per_cpu_load` . После этих расчётов мы получим корректный базовый адрес нового `GDT` .

Переменные, локальные для каждого процессора (`per-CPU variables`), являются особенностью ядра версии 2.6. Вы уже можете понять что это, исходя из названия. Когда мы создаём `per-CPU` переменную, каждый процессор будет иметь свою собственную копию этой переменной. Здесь мы создаём `per-CPU` переменную `gdt_page` . Существует много преимуществ для переменных этого типа, например, нет блокировок, поскольку каждый процессор работает со своей собственной копией переменной и т.д. Таким образом, каждое ядро на многопроцессорной машине будет иметь свою собственную таблицу `GDT` и каждая запись в таблице будет представлять сегмент памяти, к которому можно получить доступ из потока, который запускался на ядре. Подробнее о `per-CPU` переменных можно почитать в статье Concepts/linux-cpu-1.

После загрузки новой глобальной таблицы дескрипторов мы перезагружаем сегменты:

```
    xorl %eax,%eax
    movl %eax,%ds
    movl %eax,%ss
    movl %eax,%es
    movl %eax,%fs
    movl %eax,%gs
```

После всех этих шагов мы настраиваем регистр `gs` , указывающий на `irqstack` , который представляет собой специальный стек для обработки прерываний:

```
    movl    $MSR_GS_BASE,%ecx
    movl    initial_gs(%rip),%eax
    movl    initial_gs+4(%rip),%edx
    wrmsr
```

где `MSR_GS_BASE` :

```
#define MSR_GS_BASE             0xc0000101
```

Нам необходимо поместить `MSR_GS_BASE` в регистр `ecx` и загрузить данные из `eax` и `edx` (которые указывают на `initial_gs`) с помощью инструкции `wrmsr` . Мы не используем регистры сегментов `cs` , `fs` , `ds` и `ss` для адресации в 64-битном режиме, но могут использоваться регистры `fs` и `gs` . `fs` и `gs` имеют скрытую часть (как мы видели в режиме реальных адресов для `cs`) и эта часть содержит дескриптор, который отображён на моделезависимый регистр. Таким образом, выше мы можем видеть `0xc0000101` - это MSR-адрес `gs.base`. В точке входа нет стека ядра, поэтому когда происходит системный вызов или прерывание, значение `MSR_GS_BASE` будет хранить адрес стека прерываний.

На следующем шаге мы помещаем адрес структуры параметров загрузки режима реальных адресов в регистр `rdi` (напомним, что `rsi` содержит указатель на эту структуру с самого начала) и переходим к коду на C:

```
    pushq    $.Lafter_lret    # поиещает адрес возврата в стек для unwinder'a
    xorq    %rbp, %rbp    # очищает указатель фрейма
    movq    initial_code(%rip), %rax
    pushq    $__KERNEL_CS    # устанавливает корректный cs
    pushq    %rax        # целевой адрес в отрицательном пространстве
    lretq
.Lafter_lret:
```

Здесь мы помещаем адрес `initial_code` в `rax` и помещаем возвращаемый адрес `__KERNEL_CS` и адрес `initial_code` в стек. После этого мы видим инструкцию `lretq`, означающую что после неё адрес возврата будет извлечён из стека (теперь это адрес `initial_code`) и будет совершён переход по нему. `initial_code` определён в том же файле исходного кода и выглядит следующим образом:

```
    .balign    8
    GLOBAL(initial_code)
    .quad    x86_64_start_kernel
    ...
    ...
    ...
```

Как мы видим `initial_code` содержит адрес `x86_64_start_kernel`, определённой в arch/x86/kerne/head64.c:

```
asmlinkage __visible void __init x86_64_start_kernel(char * real_mode_data)
{
    ...
    ...
    ...
}
```

У неё есть один аргумент - `real_mode_data` (помните, ранее мы помещали адрес данных режима реальных адресов в регистр `rdi`).

Далее в start_kernel

Мы увидим последние приготовления, прежде чем сможем перейти к "точке входа в ядро" - к функции `start_kernel` в файле init/main.c.

Прежде всего в функции `x86_64_start_kernel` мы видим некоторый проверки:

```
BUILD_BUG_ON(MODULES_VADDR < __START_KERNEL_map);
BUILD_BUG_ON(MODULES_VADDR - __START_KERNEL_map < KERNEL_IMAGE_SIZE);
BUILD_BUG_ON(MODULES_LEN + KERNEL_IMAGE_SIZE > 2*PUD_SIZE);
BUILD_BUG_ON((__START_KERNEL_map & ~PMD_MASK) != 0);
BUILD_BUG_ON((MODULES_VADDR & ~PMD_MASK) != 0);
BUILD_BUG_ON(!(MODULES_VADDR > __START_KERNEL));
MAYBE_BUILD_BUG_ON(!(((MODULES_END - 1) & PGDIR_MASK) == (__START_KERNEL & PGDIR_MASK)));
BUILD_BUG_ON(__fix_to_virt(__end_of_fixed_addresses) <= MODULES_END);
```

например, виртуальный адрес пространства модуля не меньше, чем базовый адрес кода ядра (`__STAT_KERNEL_map`), код ядра с модулями не меньше образа ядра и т.д. `BUILD_BUG_ON` является макросом и выглядит следующим образом:

```
#define BUILD_BUG_ON(condition) ((void)sizeof(char[1 - 2*!!(condition)]))
```

Давайте попробуем понять, как работает этот трюк. Возьмём, например, первое условие: `MODULES_VADDR < __START_KERNEL_map`. `!!conditions` тоже самое что и `condition != 0`. Таким образом, если `MODULES_VADDR < __START_KERNEL_map` истинно, мы получим `1` в `!!(condition)` или ноль, если ложно. После `2*!!(condition)` мы получим или `2` или `0`. В конце вычислений мы можем получить два разных поведения:

- У нас будет ошибка компиляции, поскольку мы попытаемся получить размер `char` массива с отрицательным индексом (вполне возможно, но в нашем случае `MODULES_VADDR` не может быть меньше `__START_KERNEL_map`);
- Ошибки компиляции не будет.

На этом всё. Очень интересный C-трюк для получения ошибки компиляции, которая зависит от некоторых констант.

На следующем шаге мы видим вызов функции `cr4_init_shadow`, которая сохраняет копии `cr4` для каждого процессора. Переключения контекста могут изменять биты в `cr4`, поэтому нам нужно сохранить `cr4` для каждого процессора. После этого происходит вызов функции `reset_early_page_tables`, которая сбрасывает все записи глобального каталога страниц и записывает новый указатель на PGT в `cr3`:

```
    memset(early_top_pgt, 0, sizeof(pgd_t)*(PTRS_PER_PGD-1));
    next_early_pgt = 0;
    write_cr3(__sme_pa_nodebug(early_top_pgt));
```

Вскоре мы создадим новые таблицы страниц. Далее в цикле мы обнуляем все записи глобального каталога страниц. После этого мы устанавливаем `next_early_pgt` в ноль (подробнее об этом в следующей статье) и записываем физический адрес `early_top_pgt` в `cr3`.

После этого мы очищаем `_bss` от `__bss_stop` до `__bss_start`, а также `init_top_pgt`. `init_top_pgt` определён в arch/x86/kerne/head_64.S:

```
NEXT_PGD_PAGE(init_top_pgt)
    .fill    512,8,0
    .fill    PTI_USER_PGD_FILL,8,0
```

Это то же самое определение, что и `early_top_pgt`.

Следующим шагом будет настройка начальных обработчиков `IDT`. Это большой раздел, поэтому мы увидим его в следующей статье.

Заключение

Это конец первой части об инициализации ядра Linux.

В следующей части мы увидим инициализацию начальных обработчиков прерываний, отображение памяти пространства ядра и многое другое.

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- Моделезависимый регистр
- Страничная организация памяти

- Предыдущая часть - Рандомизация адреса ядра
- Бит NX
- ASLR

Инициализация ядра. Часть 2.

Начальная обработка прерываний и исключений

В предыдущей части мы остановились перед настройкой начальных обработчиков прерываний. На данный момент мы находимся в распакованном ядре Linux, у нас есть базовая структура страничной организации памяти для начальной загрузки, и наша текущая цель - завершить начальную подготовку до того, как основной код ядра начнёт свою работу.

Мы уже начали эту подготовку в предыдущей первой части этой главы. Мы продолжим в этой части и узнаем больше об обработке прерываний и исключений.

Как вы можете помнить, мы остановились перед этим циклом:

```
for (i = 0; i < NUM_EXCEPTION_VECTORS; i++)
    set_intr_gate(i, early_idt_handler_array[i]);
```

из файла arch/x86/kernel/head64.c. Но прежде чем начать разбирать этот код, нам нужно знать о прерываниях и обработчиках.

Некоторая теория

Прерывание - это событие, вызванное программным или аппаратным обеспечением в CPU. Например, пользователь нажал клавишу на клавиатуре. Во время прерывания, CPU останавливает текущую задачу и передаёт управление специальной процедуре - обработчику прерываний. Обработчик прерываний обрабатывает прерывания и передаёт управление обратно к ранее остановленной задаче. Мы можем разделить прерывания на три типа:

- Программные прерывания - когда программное обеспечение сигнализирует CPU, что ему нужно обратиться к ядру. Эти прерывания обычно используются для системных вызовов;
- Аппаратные прерывания - когда происходит аппаратное событие, например нажатие кнопки на клавиатуре;
- Исключения - прерывания, генерируемые процессором, когда CPU обнаруживает ошибку, например деление на ноль или доступ к странице памяти, которая не находится в ОЗУ.

Каждому прерыванию и исключению присваивается уникальный номер - `номер вектора`. `Номер вектора` может быть любым числом от `0` до `255`. Существует обычная практика использовать первые `32` векторных номеров для исключений, а номера от `32` до `255` для пользовательских прерываний. Мы можем видеть это в коде выше - `NUM_EXCEPTION_VECTORS`, определённый как:

```
#define NUM_EXCEPTION_VECTORS 32
```

CPU использует номер вектора как индекс в `таблице векторов прерываний` (мы рассмотрим её позже). Для перехвата прерываний CPU использует APIC. В следующей таблице показаны исключения `0-31`:

```
----------------------------------------------------------------------------------------------------
|Вектор|Мнемоника|Описание              |Тип     |Код ошибки|Источник                                 |
----------------------------------------------------------------------------------------------------
|0     | #DE     |Деление на ноль       |Ошибка  |Нет       |DIV и IDIV                               |
|---------------------------------------------------------------------------------------------------
|1     | #DB     |Зарезервировано       |О/Л     |Нет       |                                         |
|---------------------------------------------------------------------------------------------------
|2     | ---     |Немаск. прерывания    |Прерыв.|Нет       |Внешние NMI                               |
|---------------------------------------------------------------------------------------------------
|3     | #BP     |Исключение отладки    |Ловушка|Нет       |INT 3                                     |
```

```
|----------------------------------------------------------------------------------------------------------
|4     | #OF     |Переполнение          |Ловушка|Нет       |Инструкция INTO                        |
|----------------------------------------------------------------------------------------------------------
|5     | #BR     |Выход за границы      |Ошибка |Нет       |Инструкция BOUND                       |
|----------------------------------------------------------------------------------------------------------
|6     | #UD     |Неверный опкод        |Ошибка |Нет       |Инструкция UD2                         |
|----------------------------------------------------------------------------------------------------------
|7     | #NM     |Устройство недоступно |Ошибка |Нет       |Плавающая точка или [F]WAIT            |
|----------------------------------------------------------------------------------------------------------
|8     | #DF     |Двойная ошибка        |Авария |Да        |Инструкция, которую могут генерировать NMI |
|----------------------------------------------------------------------------------------------------------
|9     | ---     |Зарезервировано       |Ошибка |Нет       |                                       |
|----------------------------------------------------------------------------------------------------------
|10    | #TS     |Неверный TSS          |Ошибка |Да        |Смена задачи или доступ к TSS          |
|----------------------------------------------------------------------------------------------------------
|11    | #NP     |Сегмент отсутствует   |Ошибка |Нет       |Доступ к регистру сегмента             |
|----------------------------------------------------------------------------------------------------------
|12    | #SS     |Ошибка сегмента стека |Ошибка |Да        |Операции со стеком                     |
|----------------------------------------------------------------------------------------------------------
|13    | #GP     |Общее нарушение защиты|Ошибка |Да        |Ссылка на память                       |
|----------------------------------------------------------------------------------------------------------
|14    | #PF     |Ошибка страницы       |Ошибка |Да        |Ссылка на память                       |
|----------------------------------------------------------------------------------------------------------
|15    | ---     |Зарезервировано       |       |Нет       |                                       |
|----------------------------------------------------------------------------------------------------------
|16    | #MF     |Ошибка x87 FPU        |Ошибка |Нет       |Плавающая точка или [F]WAIT            |
|----------------------------------------------------------------------------------------------------------
|17    | #AC     |Проверка выравнивания |Ошибка |Да        |Ссылка на данные                       |
|----------------------------------------------------------------------------------------------------------
|18    | #MC     |Проверка машины       |Авария |Нет       |                                       |
|----------------------------------------------------------------------------------------------------------
|19    | #XM     |Исключение SIMD       |Ошибка |Нет       |Инструкции SSE[2,3]                    |
|----------------------------------------------------------------------------------------------------------
|20    | #VE     |Искл. виртуализации   |Ошибка |Нет       |Гипервизор                             |
|----------------------------------------------------------------------------------------------------------
|21-31 | ---     |Зарезервировано       |Прерыв.|Нет       |Внешние прерывания                     |
 ----------------------------------------------------------------------------------------------------------
```

Исключения делятся на три типа:

- Ошибки (Faults) - исключения, по окончании обработки которых прерванная команда повторяется;
- Ловушки (Traps) - исключения, при обработке которых CPU сохраняет состояние, следующее за командой, вызвавшей исключение;
- Аварии (Aborts) - исключения, при обработке которых CPU не сохраняет состояния и не имеет возможности вернуться к месту исключения

Для реагирования на прерывание CPU использует специальную структуру - таблицу векторов прерываний (Interrupt Descriptor Table, IDT). IDT является массивом 8-байтных дескрипторов, наподобие глобальной таблицы дескрипторов, но записи в IDT называются `шлюзами` (gates). CPU умножает номер вектора на 8 для того чтобы найти индекс записи IDT. Но в 64-битном режиме IDT представляет собой массив 16-байтных дескрипторов и CPU умножает номер вектора на 16. Из предыдущей части мы помним, что CPU использует специальный регистр `GDTR` для поиска глобальной таблицы дескрипторов, поэтому CPU использует специальный регистр `IDTR` для таблицы векторов прерываний и инструкцию `lidt` для загрузки базового адреса таблицы в этот регистр.

Запись IDT в 64-битном режиме имеет следующую структуру:

```
127                                                                             96
 --------------------------------------------------------------------------------
|                                                                               |
|                                Зарезервировано                                |
|                                                                               |
 --------------------------------------------------------------------------------
95                                                                              64
 --------------------------------------------------------------------------------
```

```
|                                                                               |
|                               Смещение 63..32                                 |
|                                                                               |
 --------------------------------------------------------------------------------
63                               48 47      46  44   42    39            34    32
 --------------------------------------------------------------------------------
|                                  |       |  D  |   |     |      |   |   |     |
|       Смещение 31..16            |   P   |  P  | 0 |Тип  |0 0 0 | 0 | 0 | IST |
|                                  |       |  L  |   |     |      |   |   |     |
 --------------------------------------------------------------------------------
31                                   16 15                                      0
 --------------------------------------------------------------------------------
|                                      |                                        |
|          Селектор сегмента           |                 Смещение 15..0         |
|                                      |                                        |
 --------------------------------------------------------------------------------
```

где:

- `Смещение` - смещение к точки входа обработчика прерывания;
- `DPL` - уровень привилегий сегмента (Descriptor Privilege Level);
- `P` - флаг присутствия сегмента;
- `Селектор сегмента` - селектор сегмента кода в GDT или LDT
- `IST` - обеспечивает возможность переключения на новый стек для обработки прерываний.

И последнее поле `Тип` описывает тип записи `IDT`. Существует три различных типа обработчиков для прерываний:

- Дескриптор задачи
- Дескриптор прерывания
- Дескриптор ловушки

Дескрипторы прерываний и ловушек содержат дальний указатель на точку входа обработчика прерываний. Различие между этими типами заключается в том, как CPU обрабатывает флаг `IF`. Если обработчик прерываний был вызван через шлюз прерывания, CPU очищает флаг `IF` чтобы предотвратить другие прерывания, пока выполняется текущий обработчик прерываний. После выполнения текущего обработчика прерываний CPU снова устанавливает флаг `IF` с помощью инструкции `iret`.

Остальные биты в шлюзе прерывания зарезервированы и должны быть равны 0. Теперь давайте посмотрим, как CPU обрабатывает прерывания:

- CPU сохраняет регистр флагов, `CS`, и указатель на инструкцию в стеке.
- Если прерывание вызывает код ошибки (например, `#PF`), CPU сохраняет ошибку в стеке после указателя на инструкцию;
- После выполнения обработчика прерываний для возврата из него используется инструкция `iret`.

Теперь вернёмся к коду.

Заполнение и загрузка IDT

Мы остановились на следующем моменте:

```
for (i = 0; i < NUM_EXCEPTION_VECTORS; i++)
    set_intr_gate(i, early_idt_handler_array[i]);
```

Здесь мы вызываем `set_intr_gate` в цикле, который принимает два параметра:

- Номер прерывания или `номер вектора`;
- Адрес обработчика idt.

и вставляет шлюз прерывания в таблицу `IDT` , которая представлена массивом `&idt_descr` . Прежде всего, давайте посмотрим на массив `early_idt_handler_array` . Это массив, который определён в заголовочном файле arch/x86/include/asm/segment.h и содержит адреса первых `32` обработчиков исключений:

```
#define EARLY_IDT_HANDLER_SIZE   9
#define NUM_EXCEPTION_VECTORS    32

extern const char early_idt_handler_array[NUM_EXCEPTION_VECTORS][EARLY_IDT_HANDLER_SIZE];
```

The `early_idt_handler_array` - это `288` байтный массив, который содержит адреса точек входа обработчиков исключений каждые девять байт. Каждый девять байт этого массива состоят из двух байт необязательной инструкции для помещения фиктивного кода ошибки, если исключение не предоставляет его, двубайтовая инструкция для помещения номера вектора в стек и пять байт `jump` на общий код обработчика исключений.

Как можно видеть, в цикле мы заполняем только первые 32 элемента `IDT` , поскольку все начальные настройки запускаются с отключёнными прерываниями, поэтому нет необходимости настраивать обработчики прерываний для векторов, превышающих `32` . В массиве `early_idt_handler_array` содержатся общий обработчики idt и мы можем найти его определение в ассемблерном файле arch/x86/kernel/head_64.S. Пока что мы пропустим его, но вскоре вернёмся к нему. Перед этим мы рассмотрим реализацию функции `set_intr_gate` .

Функция `set_intr_gate` определена в файле arch/x86/kernel/idt.c:

```
static void set_intr_gate(unsigned int n, const void *addr)
{
    struct idt_data data;

    BUG_ON(n > 0xFF);

    memset(&data, 0, sizeof(data));
    data.vector     = n;
    data.addr     = addr;
    data.segment     = __KERNEL_CS;
    data.bits.type     = GATE_INTERRUPT;
    data.bits.p     = 1;

        idt_setup_from_table(idt_table, &data, 1, false);
}
```

Прежде всего она проверяет, что переданный номер прерывания не больше чем `255` с помощью макроса `BUG_ON` . Нам нужно сделать эту проверку, поскольку максимально возможное количество прерываний - `256` . Далее мы устаналиваем данные IDT заданными значениями. И уже после этого мы вызываем функцию `idt_setup_from_table` :

```
static void
idt_setup_from_table(gate_desc *idt, const struct idt_data *t, int size, bool sys)
{
    gate_desc desc;

    for (; size > 0; t++, size--) {
        desc.offset_low    = (u16) t->addr;
        desc.segment       = (u16) t->segment
        desc.bits       = t->bits;
        desc.offset_middle = (u16) (t->addr >> 16);
        desc.offset_high   = (u32) (t->addr >> 32);
        desc.reserved       = 0;
        memcpy(&idt[t->vector], &desc, sizeof(desc));
        if (sys)
            set_bit(t->vector, system_vectors);
    }
}
```

которая заполняет три части адреса обработчика прерываний адресом, который мы получили в основном цикле (адрес точки входа обработчика прерывания). Далее мы просто копируем дескриптор шлюза в запись IDT.

После завершения основного цикла у нас в распоряжении будет заполненный массив `idt_table` структур `gate_desc` и теперь мы можем загрузить `таблицу векторов прерываний` вызовом:

```
load_idt((const struct desc_ptr *)&idt_descr);
```

Где `idt_descr` :

```
struct desc_ptr idt_descr __ro_after_init = {
    .size        = (IDT_ENTRIES * 2 * sizeof(unsigned long)) - 1,
    .address     = (unsigned long) idt_table,
};
```

и `load_idt` просто выполняет инструкцию `lidt` :

```
asm volatile("lidt %0"::"m" (*dtr));
```

Мы можем заметить, что вызовы функций `_trace_*` есть в `_set_gate` и в остальных функциях. Эти функции заполняют шлюзы `IDT` таким же образом, что и `_set_gate` , но с одним отличием. Эти функции используют `trace_idt_table` `таблицы векторов прерываний` вместо `idt_table` для контрольных точек (мы рассмотрим эту тему в другой части).

Итак, мы заполнили и загрузили `таблицу векторов прерываний` и мы знаем как ведёт себя CPU во время прерывания. Теперь самое время перейти к обработчикам прерываний.

Начальные обработчики прерываний

Как говорилось ранее, мы заполнили `IDT` адресом `early_idt_handler_array` . Мы можем найти его в arch/x86/kernel/head_64.S:

```
ENTRY(early_idt_handler_array)
    i = 0
    .rept NUM_EXCEPTION_VECTORS
    .if ((EXCEPTION_ERRCODE_MASK >> i) & 1) == 0
        UNWIND_HINT_IRET_REGS
        pushq $0    # Dummy error code, to make stack frame uniform
    .else
        UNWIND_HINT_IRET_REGS offset=8
    .endif
    pushq $i        # 72(%rsp) Vector number
    jmp early_idt_handler_common
    UNWIND_HINT_IRET_REGS
    i = i + 1
    .fill early_idt_handler_array + i*EARLY_IDT_HANDLER_SIZE - ., 1, 0xcc
    .endr
    UNWIND_HINT_IRET_REGS offset=16
END(early_idt_handler_array)
```

Функции, которые, как правило, не вызываются напрямую другими функциями, такие как `syscall` и обработчики прерываний, часто имеют необычные вещи вроде функции не C-типа с указателем стека. Такой код необходимо аннотировать с помощью макроса `UNWIND_HINT_IRET_REGS` , чтобы `objtool` смог его понять. Здесь мы видим создание обработчиков прерываний для первых `32` исключений. Мы проверяем, содержит ли исключение код ошибки, и ничего не делаем, если исключение не возвращает код ошибки, тогда мы помещаем в стек ноль. Мы делаем это для того чтобы стек был однородным. После этого мы помещаем номер исключения в стек и переходим на `early_idt_handler_array` , который является общим обработчиком прерываний на данный момент. Каждый девятый байт массива `early_idt_handler_array` состоит из необязательного кода ошибок, `номера вектора` и инструкции перехода. Мы можем видеть это в выводе утилиты `objdump` :

```
$ objdump -D vmlinux
...
...
...
ffffffff81fe5000 <early_idt_handler_array>:
ffffffff81fe5000:       6a 00                   pushq  $0x0
ffffffff81fe5002:       6a 00                   pushq  $0x0
ffffffff81fe5004:       e9 17 01 00 00          jmpq   ffffffff81fe5120 <early_idt_handler_common>
ffffffff81fe5009:       6a 00                   pushq  $0x0
ffffffff81fe500b:       6a 01                   pushq  $0x1
ffffffff81fe500d:       e9 0e 01 00 00          jmpq   ffffffff81fe5120 <early_idt_handler_common>
ffffffff81fe5012:       6a 00                   pushq  $0x0
ffffffff81fe5014:       6a 02                   pushq  $0x2
...
...
...
```

Как я писал ранее, CPU помещает регистр флагов, `CS` и `RIP` в стек. Поэтому, прежде чем `early_idt_handler` будет выполнен, стек будет содержать следующие данные:

```
|--------------------|
| %rflags            |
| %cs                |
| %rip               |
| код ошибки         | <-- %rsp
|--------------------|
```

Давайте посмотрим на реализацию `early_idt_handler_common` . Он находится в том же ассемблерном файле arch/x86/kernel/head_64.S и прежде всего инкрементирует `early_recursion_flag` , чтобы предотвратить рекурсию в `early_idt_handler_common` :

```
    incl early_recursion_flag(%rip)
```

Далее мы сохраняем регистры общего назначения в стек:

```
    pushq %rsi
    movq 8(%rsp), %rsi
    movq %rdi, 8(%rsp)
    pushq %rdx
    pushq %rcx
    pushq %rax
    pushq %r8
    pushq %r9
    pushq %r10
    pushq %r11
    pushq %rbx
    pushq %rbp
    pushq %r12
    pushq %r13
    pushq %r14
    pushq %r15
    UNWIND_HINT_REGS
```

Мы должны сделать это, чтобы предотвратить появление неверных значений регистров при возврате из обработчика прерываний. После этого мы проверяем номер вектора, и если он `#PF` или ошибка страницы (Page Fault), мы помещаем значение регистра `cr2` в регистр `rdi` и вызываем `early_make_pgtable` (мы скоро это увидим):

```
    cmpq $14,%rsi
    jnz 10f
    GET_CR2_INTO(%rdi)
    call early_make_pgtable
    andl %eax,%eax
    jz 20f
```

Если номер вектора не равен `#PF` , мы вызываем функцию `early_fixup_exception` с передачей указателя ядра (смотрите соглашение о вызовах x86-64):

```
10:
    movq %rsp,%rdi
    call early_fixup_exception
```

Мы увидим реализацию функции `early_fixup_exception` позже.

```
20:
    decl early_recursion_flag(%rip)
    jmp restore_regs_and_return_to_kernel
```

После того как мы уменьшим значение параметра `initial_recursion_flag` , мы восстанавливаем регистры из стека, которые мы сохранили ранее, и возвращаемся из обработчика с помощью `iretq` .

Это конец первого обработчика прерываний. Обратите внимание, что это очень ранний обработчик прерываний, поэтому он обрабатывает только ошибку страницы. Мы увидим обработчики и для других прерываний, но пока давайте посмотрим на обработчик ошибки страницы.

Обработка ошибки страницы

В предыдущем разделе мы увидели первый начальный обработчик прерываний, который проверяет, что номер прерывания относится к ошибке страницы и вызывает `early_make_pgtable` для создания новых таблиц страниц. На данном этапе нам необходим обработчик `#PF` , поскольку планируется добавить способность загружать ядро выше `4G` и сделать структуру `boot_params` доступной над 4G.

Вы можете найти реализацию `early_make_pgtable` в arch/x86/kernel/head64.c и он принимает только один параметр - адрес из регистра `cr2` , который вызывал ошибку страницы. Давайте посмотрим на неё более подробно:

```
int __init early_make_pgtable(unsigned long address)
{
    unsigned long physaddr = address - __PAGE_OFFSET;
    pmdval_t pmd;

    pmd = (physaddr & PMD_MASK) + early_pmd_flags;

    return __early_make_pgtable(address, pmd);
}
```

Затем мы вызываем функцию `__early_make_pgtable` , которая определена в том же файле, что и функция `early_make_pgtable` , как показано ниже:

```
int __init __early_make_pgtable(unsigned long address, pmdval_t pmd)
{
    unsigned long physaddr = address - __PAGE_OFFSET;
    pgdval_t pgd, *pgd_p;
    p4dval_t p4d, *p4d_p;
    pudval_t pud, *pud_p;
```

```
    pmdval_t *pmd_p;
    ...
    ...
    ...
}
```

Она начинается с определения некоторых переменных, которые имеют типы `*val_t`. Все эти типы всего-навсего:

```
typedef unsigned long   pgdval_t;
```

Также мы будем работать с типами `*_t`, например `pgd_t` и т.д. Все эти типы определены в arch/x86/include/asm/pgtable_types.h и представляют собой структуры:

```
typedef struct { pgdval_t pgd; } pgd_t;
```

Для примера,

```
extern pgd_t early_top_pgt[PTRS_PER_PGD];
```

Здесь `early_top_pgt` представляет начальный каталог таблиц страниц верхнего уровня, который состоит из массива типа `pgd_t` и `pgd` указывает на записи страниц нижнего уровня.

После того как мы проверили, что у нас корректный адрес, мы получаем адрес записи глобального каталога страниц, который содержит адрес `#PF`, и присваиваем его значение переменной `pgd`:

```
pgd_p = &early_top_pgt[pgd_index(address)].pgd;
pgd = *pgd_p;
```

Затем мы проверяем, включена ли пятиуровневая страничная организация памяти:

```
if (!pgtable_l5_enabled())
    p4d_p = pgd_p;
```

В большинстве случаев пятиуровневая страничная организация не включена, поэтому `p4d_p` скорее всего равен `pgd_p`.

После этого мы исправляем адрес `p4d` с помощью:

```
p4d_p += p4d_index(address);
p4d = *p4d_p;
```

На следующем шаге мы проверяем `p4d`, если он содержит верную запись p4d, мы помещаем физический адрес записи p4d в `pud_p`:

```
pud_p = (pudval_t *)((p4d & PTE_PFN_MASK) + __START_KERNEL_map - phys_base);
```

где `PTE_PFN_MASK` является макросом:

```
#define PTE_PFN_MASK            ((pteval_t)PHYSICAL_PAGE_MASK)
```

который раскрывается до:

```
(signed long)(~(PAGE_SIZE-1)) & ((1 << 52) - 1)
```

Здесь используется расширение знака. Развёрнутая форма:

```
0b1111111111111111111111111111111111111111111111111111
```

т.е макрос состоит из 52 бит для маскирования фрейма страниц.

Если `p4d` не содержит верный адрес, мы проверяем что `next_early_pgt` не больше чем `EARLY_DYNAMIC_PAGE_TABLES`, который равен `64` и представляет фиксированное количество буферов для настройки новых таблиц страниц по требованию. Если `next_early_pgt` больше, чем `EARLY_DYNAMIC_PAGE_TABLES` мы сбрасываем таблицы страниц и начинаем всё заново. Если `next_early_pgt` меньше, чем `EARLY_DYNAMIC_PAGE_TABLES`, мы создаём новый указатель верхнего каталога страниц, который указывает на текущую динамическую таблицу страниц и записываем его физический адрес с правами доступа `_KERNPG_TABLE` в p4d:

```
if (next_early_pgt >= EARLY_DYNAMIC_PAGE_TABLES) {
    reset_early_page_tables();
    goto again;
}

pud_p = (pudval_t *)early_dynamic_pgts[next_early_pgt++];
for (i = 0; i < PTRS_PER_PUD; i++)
    pud_p[i] = 0;
*p4d_p = (p4dval_t)pud_p - __START_KERNEL_map + phys_base + _KERNPG_TABLE;
```

Как уже говорили ранее, мы исправляем адрес верхнего каталога страниц:

```
pud_p += pud_index(address);
pud = *pud_p;
```

На следующем шаге мы делаем те же действия что и ранее, но с промежуточным каталогом страниц. В конце мы исправляем адрес промежуточного каталога страниц, который содержит отображения текста ядра+виртуальные адреса данных:

```
pmd_p[pmd_index(address)] = pmd;
```

После того как обработчик ошибки страницы завершён, `early_top_pgt` содержит записи, которые указывают на корректные адреса.

Другие обработчики исключений

На раннем этапе прерывания, исключения, кроме ошибки страницы, обрабатываются функцией `early_fixup_exception`, которая определена в arch/x86/mm/extable.c и принимает два аргумента - указатель на стек ядра, состоящий из сохраненных регистров, и номер прерывания:

```
void __init early_fixup_exception(struct pt_regs *regs, int trapnr)
{
    ...
    ...
    ...
}
```

Прежде всего нам нужно проверить некоторые условия:

```
    if (trapnr == X86_TRAP_NMI)
        return;
```

```
    if (early_recursion_flag > 2)
        goto halt_loop;
    if (!xen_pv_domain() && regs->cs != __KERNEL_CS)
        goto fail;
```

Здесь мы просто игнорируем NMI. Также мы убеждаемся, что мы не в рекурсивной ситуации.

Далее:

```
    if (fixup_exception(regs, trapnr))
        return;
```

Функция `fixup_exception` определена в том же файле, что и `early_fixup_exception` :

```
int fixup_exception(struct pt_regs *regs, int trapnr)
{
    const struct exception_table_entry *e;
    ex_handler_t handler;
    e = search_exception_tables(regs->ip);
    if (!e)
        return 0;
    handler = ex_fixup_handler(e);
    return handler(e, regs, trapnr);
}
```

`ex_handler_t` - тип указателя функции, определённый следующий образом:

```
typedef bool (*ex_handler_t)(const struct exception_table_entry *,
                            struct pt_regs *, int)
```

Функция `search_exception_tables` ищет данный адрес в таблице исключений (т.е содержимое ELF-секции `__ex_table`). После этого мы получаем фактический адрес с помощью функции `ex_fixup_handler` . Наконец, мы вызываем фактический обработчик. Для получения дополнительной информации о таблице исключений смотрите Documentation/x86/exception-tables.txt.

Вернёмся к функции `early_fixup_exception` . Следующий шаг:

```
    if (fixup_bug(regs, trapnr))
        return;
```

Функция `fixup_bug` определена в arch/x86/kernel/traps.c. Давайте посмотрим на реализацию этой функции.

```
int fixup_bug(struct pt_regs *regs, int trapnr)
{
    if (trapnr != X86_TRAP_UD)
        return 0;
    switch (report_bug(regs->ip, regs)) {
    case BUG_TRAP_TYPE_NONE:
    case BUG_TRAP_TYPE_BUG:
        break;
    case BUG_TRAP_TYPE_WARN:
        regs->ip += LEN_UD2;
        return 1;
    }
    return 0;
}
```

Всё что делает эта функция - возвращает `1` , если генерируется исключение `#UD` (или неверный опкод (Invalid Opcode)) и функция `report_bug` вернула `BUG_TRAP_TYPE_WARN` , иначе возвращает `0` .

Заключение

Это конец второй части инициализации ядра Linux. В следующей части мы увидим все шаги перед точкой входа в ядро - функции `start_kernel` .

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- GNU assembly .rept
- APIC
- NMI
- Таблица страниц
- Обработчик прерываний
- Ошибка страницы
- Предыдущая часть

Инициализация ядра. Часть 3.

Последние приготовления перед точкой входа в ядро

Это третья часть серии Инициализация ядра. В предыдущей части мы увидели начальную обработку прерываний и исключений и продолжим погружение в процесс инициализации ядра Linux в текущей части. Наша следующая точка - "точка входа в ядро" - функция `start_kernel` из файла init/main.c. Да, технически это не точка входа в ядро, а начало кода ядра, который не зависит от определённой архитектуры. Но прежде чем мы вызовем функцию `start_kernel`, мы должны совершить некоторые приготовления. Давайте продолжим.

Снова boot_params

В предыдущей части мы остановились на настройке таблицы векторов прерываний и её загрузки в регистр `IDTR`. На следующем шаге мы можем видеть вызов функции `copy_bootdata`:

```
copy_bootdata(__va(real_mode_data));
```

Эта функция принимает один аргумент - виртуальный адрес `real_mode_data`. Вы должны помнить, что мы передали адрес структуры `boot_params` из arch/x86/include/uapi/asm/bootparam.h в функцию `x86_64_start_kernel` как первый параметр в arch/x86/kernel/head_64.S:

```
	/* rsi is pointer to real mode structure with interesting info.
	   pass it to C */
	movq	%rsi, %rdi
```

Взглянем на макрос `__va`. Этот макрос определён в init/main.c:

```
#define __va(x)                 ((void *)((unsigned long)(x)+PAGE_OFFSET))
```

где `PAGE_OFFSET` это `__PAGE_OFFSET` (`0xffff880000000000` и базовый виртуальный адрес прямого отображения всей физической памяти). Таким образом, мы получаем виртуальный адрес структуры `boot_params` и передаём его функции `copy_bootdata`, в которой мы копируем `real_mod_data` в `boot_params`, объявленный в файле arch/x86/include/asm/setup.h

```
extern struct boot_params boot_params;
```

Давайте посмотрим на реализацию `copy_boot_data`:

```
static void __init copy_bootdata(char *real_mode_data)
{
	char * command_line;
	unsigned long cmd_line_ptr;

	memcpy(&boot_params, real_mode_data, sizeof boot_params);
	sanitize_boot_params(&boot_params);
	cmd_line_ptr = get_cmd_line_ptr();
	if (cmd_line_ptr) {
		command_line = __va(cmd_line_ptr);
		memcpy(boot_command_line, command_line, COMMAND_LINE_SIZE);
	}
}
```

Прежде всего, обратите внимание на то, что эта функция объявлена с префиксом `__init`. Это означает, что эта функция будет использоваться только во время инициализации и используемая память будет освобождена.

Мы можем видеть объявление двух переменных для командной строки ядра и копирование `real_mode_data` в `boot_params` функцией `memcpy`. Далее следует вызов функции `sanitize_boot_params`, которая заполняет некоторые поля структуры `boot_params`, такие как `ext_ramdisk_image` и т.д, если загрузчики не инициализировал неизвестные поля в `boot_params` нулём. После этого мы получаем адрес командной строки вызовом функции `get_cmd_line_ptr`:

```
unsigned long cmd_line_ptr = boot_params.hdr.cmd_line_ptr;
cmd_line_ptr |= (u64)boot_params.ext_cmd_line_ptr << 32;
return cmd_line_ptr;
```

который получает 64-битный адрес командной строки из заголовочного файла загрузки ядра и возвращает его. На последнем шаге мы проверяем `cmd_line_ptr`, получаем его виртуальный адрес и копируем его в `boot_command_line`, который представляет собой всего лишь массив байтов:

```
extern char __initdata boot_command_line[];
```

После этого мы имеем скопированную командную строку ядра и структуру `boot_params`. На следующем шаге происходит вызов функции `load_ucode_bsp`, которая загружает процессорный микрокод, его мы здесь не увидим.

После загрузки микрокода мы можем видеть проверку функции `console_loglevel` и `early_printk`, которая печатает строку `Kernel Alive`. Но вы никогда не увидите этот вывод, потому что `early_printk` еще не инициализирован. Это небольшая ошибка в ядре, и я (*0xAX*, *автор оригинальной книги - Прим. пер.*) отправил патч - коммит, чтобы исправить её.

Перемещение по страницам инициализации

На следующем шаге, когда мы скопировали структуру `boot_params`, нам нужно перейти от начальных таблиц страниц к таблицам страниц для процесса инициализации. Мы уже настроили начальные таблицы страниц, вы можете прочитать об этом в предыдущей части и сбросили это всё функцией `reset_early_page_tables` (вы тоже можете прочитать об этом в предыдущей части) и сохранили только отображение страниц ядра. После этого мы вызываем функцию `clear_page`:

```
    clear_page(init_level4_pgt);
```

с аргументом `init_level4_pgt`, который определён в файле arch/x86/kernel/head_64.S и выглядит следующим образом:

```
NEXT_PAGE(init_level4_pgt)
    .quad   level3_ident_pgt - __START_KERNEL_map + _KERNPG_TABLE
    .org    init_level4_pgt + L4_PAGE_OFFSET*8, 0
    .quad   level3_ident_pgt - __START_KERNEL_map + _KERNPG_TABLE
    .org    init_level4_pgt + L4_START_KERNEL*8, 0
    .quad   level3_kernel_pgt - __START_KERNEL_map + _PAGE_TABLE
```

Он отображает первые 2 гигабайта и 512 мегабайта для кода ядра, данных и bss. Функция `clear_page` определена в arch/x86/lib/clear_page_64.S. Давайте взглянем на неё:

```
ENTRY(clear_page)
    CFI_STARTPROC
    xorl %eax,%eax
    movl $4096/64,%ecx
    .p2align 4
    .Lloop:
```

```
    decl    %ecx
#define PUT(x) movq %rax,x*8(%rdi)
    movq %rax,(%rdi)
    PUT(1)
    PUT(2)
    PUT(3)
    PUT(4)
    PUT(5)
    PUT(6)
    PUT(7)
    leaq 64(%rdi),%rdi
    jnz    .Lloop
    nop
    ret
    CFI_ENDPROC
    .Lclear_page_end:
    ENDPROC(clear_page)
```

Как вы можете понять из имени функции, она очищает или заполняет нулями таблицы страниц. Прежде всего обратите внимание, что эта функция начинается с макросов `CFI_STARTPROC` и `CFI_ENDPROC` , которые раскрываются до директив сборки GNU:

```
#define CFI_STARTPROC           .cfi_startproc
#define CFI_ENDPROC             .cfi_endproc
```

и используются для отладки. После макроса `CFI_STARTPROC` мы обнуляем регистр `eax` и помещаем 64 в `ecx` (это будет счётчик). Далее мы видим цикл, который начинается с метки `.Lloop` и декремента `ecx` . После этого мы помещаем нуль из регистра `rax` в `rdi` , который теперь содержит базовый адрес `init_level4_pgt` и выполняем ту же процедуру семь раз, но каждый раз перемещаем смещение `rdi` на 8. После этого первые 64 байта `init_level4_pgt` будут заполнены нулями. На следующем шаге мы снова помещаем адрес `init_level4_pgt` со смещением 64 байта в `rdi` и повторяем все операции до тех пор, пока `ecx` не будет равен нулю. В итоге мы получим `init_level4_pgt` , заполненный нулями.

После заполнения нулями `init_level4_pgt` , мы помещаем последнюю запись в `init_level4_pgt` :

```
init_level4_pgt[511] = early_top_pgt[511];
```

Вы должны помнить, что мы очистили все записи `early_top_pgt` функцией `reset_early_page_table` и сохранили только отображение ядра.

Последний шаг в функции `x86_64_start_kernel` заключается в вызове функции `x86_64_start_reservations` :

```
x86_64_start_reservations(real_mode_data);
```

с аргументов `real_mode_data` . Функция `x86_64_start_reservations` определена в том же файле исходного кода что и `x86_64_start_kernel` :

```
void __init x86_64_start_reservations(char *real_mode_data)
{
    if (!boot_params.hdr.version)
        copy_bootdata(__va(real_mode_data));

    reserve_ebda_region();

    start_kernel();
}
```

Это последняя функция перед входом в точку ядра - `start_kernel`. Давайте посмотрим, что он делает и как это работает.

Последний шаг перед точкой входа в ядро

В первую очередь мы видим проверку `boot_params.hdr.version` в функции `x86_64_start_reservations`:

```
if (!boot_params.hdr.version)
    copy_bootdata(__va(real_mode_data));
```

и если он равен нулю то снова вызывается функция `copy_bootdata` с виртуальным адресом `real_mode_data`.

В следующем шаге мы видим вызов функции `reserve_ebda_region`, определённой в файле arch/x86/kernel/head.c. Эта функция резервирует блок памяти для `EBDA` или `Extended BIOS Data Area`. `Extended BIOS Data Area` расположена в верхних адресах основной области памяти (conventional memory) и содержит данные о портах, параметрах диска и т.д.

Давайте посмотрим на функцию `reserve_ebda_region`. Он начинается с проверки, включена ли паравиртуализация или нет:

```
if (paravirt_enabled())
    return;
```

если паравиртуализация включена, мы выходим из функции `reserve_ebda_region`, потому что `EBDA` отсутствует. На следующем шаге нам нужно получить конец нижней области памяти:

```
lowmem = *(unsigned short *)__va(BIOS_LOWMEM_KILOBYTES);
lowmem <<= 10;
```

Мы получаем виртуальный адрес нижней области памяти BIOS в килобайтах и преобразуем его в байты, сдвигая его на 10 (другими словами умножаем на 1024). После этого нам нужно получить адрес `EBDA`:

```
ebda_addr = get_bios_ebda();
```

Функция `get_bios_ebda` определена в файле arch/x86/include/asm/bios_ebda.h:

```
static inline unsigned int get_bios_ebda(void)
{
    unsigned int address = *(unsigned short *)phys_to_virt(0x40E);
    address <<= 4;
    return address;
}
```

Давайте попробуем понять, как это работает. Мы видим преобразование физического адреса `0x40E` в виртуальный, где `0x0040:0x000e` - это сегмент, который содержит базовый адрес `EBDA`. Не беспокойтесь о том, что мы используем функцию `phys_to_virt` для преобразования физического адреса в виртуальный. Вы можете заметить, что ранее мы использовали макрос `__va`, но `phys_to_virt` - это то же самое:

```
static inline void *phys_to_virt(phys_addr_t address)
{
         return __va(address);
}
```

только с одним отличием: мы передаем аргумент `phys_addr_t`, который зависит от `CONFIG_PHYS_ADDR_T_64BIT`:

```
#ifdef CONFIG_PHYS_ADDR_T_64BIT
    typedef u64 phys_addr_t;
#else
    typedef u32 phys_addr_t;
#endif
```

Мы получили виртуальный адрес сегмента, в котором хранится базовый адрес `EBDA`. Мы сдвигаем его на 4 и возвращаем как результат. После этого переменная `ebda_addr` содержит базовый адрес `EBDA`.

На следующем шаге мы проверяем, что адрес `EBDA` и нижняя область памяти не меньше, чем значение макроса `INSANE_CUTOFF`:

```
if (ebda_addr < INSANE_CUTOFF)
    ebda_addr = LOWMEM_CAP;

if (lowmem < INSANE_CUTOFF)
    lowmem = LOWMEM_CAP;
```

где `INSANE_CUTOFF`:

```
#define INSANE_CUTOFF        0x20000U
```

или 128 килобайт. На последнем шаге мы получаем нижнюю часть нижней области памяти и `EBDA` и вызываем функцию `memblock_reserve`, которая резервирует область памяти для `EBDA` между нижней областью памяти и одномегабайтной меткой:

```
lowmem = min(lowmem, ebda_addr);
lowmem = min(lowmem, LOWMEM_CAP);
memblock_reserve(lowmem, 0x100000 - lowmem);
```

функция `memblock_reserve` определена в mm/block.c и принимает два аргумента:

- базовый физический адрес;
- размер области памяти.

и резервирует область памяти для заданного базового адреса и размера. `memblock_reserve` - первая функция в этой книге из фреймворка менеджера памяти ядра Linux. Мы скоро рассмотрим менеджер памяти, но пока что посмотрим на его реализацию.

Первое знакомство с фреймворком менеджера памяти ядра Linux

В предыдущем абзаце мы остановились на вызове функции `memblock_reserve` и, как я уже сказал, это первая функция из фреймворка менеджера памяти. Давайте попробуем понять, как это работает. `memblock_reserve` просто вызывает функцию:

```
memblock_reserve_region(base, size, MAX_NUMNODES, 0);
```

и передаёт ей 4 аргумента:

- физический базовый адрес области памяти;
- размер области памяти;
- максимально число NUMA-узлов;
- флаги.

В начале тела функции `memblock_reserve_region` мы можем видеть определение структуры `memblock_type` :

```
struct memblock_type *_rgn = &memblock.reserved;
```

которая представляет тип блока памяти:

```
struct memblock_type {
         unsigned long cnt;
         unsigned long max;
         phys_addr_t total_size;
         struct memblock_region *regions;
};
```

Поскольку нам необходимо зарезервировать блок памяти для `EBDA` , тип текущей области памяти зарезервирован так же, где и структура `memblock` :

```
struct memblock {
         bool bottom_up;
         phys_addr_t current_limit;
         struct memblock_type memory;
         struct memblock_type reserved;
#ifdef CONFIG_HAVE_MEMBLOCK_PHYS_MAP
         struct memblock_type physmem;
#endif
};
```

и описывает общий блок памяти. Мы инициализируем `_rgn` адресом `memblock.reserved` . `memblock` - глобальная переменная:

```
struct memblock memblock __initdata_memblock = {
    .memory.regions        = memblock_memory_init_regions,
    .memory.cnt        = 1,
    .memory.max        = INIT_MEMBLOCK_REGIONS,
    .reserved.regions    = memblock_reserved_init_regions,
    .reserved.cnt        = 1,
    .reserved.max        = INIT_MEMBLOCK_REGIONS,
#ifdef CONFIG_HAVE_MEMBLOCK_PHYS_MAP
    .physmem.regions    = memblock_physmem_init_regions,
    .physmem.cnt        = 1,
    .physmem.max        = INIT_PHYSMEM_REGIONS,
#endif
    .bottom_up        = false,
    .current_limit        = MEMBLOCK_ALLOC_ANYWHERE,
};
```

Мы не будем погружаться в детали этой переменной, но мы увидим все подробности об этом в частях о менеджере памяти. Просто отметьте, что переменная `memblock` определена с помощью `__initdata_memblock` :

```
#define __initdata_memblock __meminitdata
```

где `__meminit_data` :

```
#define __meminitdata    __section(.meminit.data)
```

Из этого можно сделать вывод, что все блоки памяти будут в секции `.meminit.data` . После того как мы определили `_rgn` , мы печатаем информацию об этом с помощью макроса `memblock_dbg` . Вы можете включить его, передав `memblock = debug` в командную строку ядра.

После печати строк отладки следует вызов функции `memblock_add_range`:

```
memblock_add_range(_rgn, base, size, nid, flags);
```

которая добавляет новую область блока памяти в секцию `.meminit.data`. Поскольку мы не инициализируем `_rgn` и он содержит `&memblock.reserved`, мы просто заполняем переданный `_rgn` базовым адресом `EBDA`, размером этой области и флагами:

```
if (type->regions[0].size == 0) {
    WARN_ON(type->cnt != 1 || type->total_size);
    type->regions[0].base = base;
    type->regions[0].size = size;
    type->regions[0].flags = flags;
    memblock_set_region_node(&type->regions[0], nid);
    type->total_size = size;
    return 0;
}
```

После заполнения нашей области памяти мы видим вызов функции `memblock_set_region_node` с двумя аргументами:

- адрес заполненной области памяти;
- id NUMA-узла.

где наши области памяти представлены структурой `memblock_region`:

```
struct memblock_region {
    phys_addr_t base;
    phys_addr_t size;
    unsigned long flags;
#ifdef CONFIG_HAVE_MEMBLOCK_NODE_MAP
    int nid;
#endif
};
```

Id NUMA-узла зависит от макроса `MAX_NUMNODES`, определённого в файле include/linux/numa.h:

```
#define MAX_NUMNODES    (1 << NODES_SHIFT)
```

где `NODES_SHIFT` зависит от параметра конфигурации `CONFIG_NODES_SHIFT`:

```
#ifdef CONFIG_NODES_SHIFT
  #define NODES_SHIFT     CONFIG_NODES_SHIFT
#else
  #define NODES_SHIFT     0
#endif
```

Функция `memblick_set_region_node` просто заполняет поле `nid` из `memblock_region` заданным значением:

```
static inline void memblock_set_region_node(struct memblock_region *r, int nid)
{
         r->nid = nid;
}
```

После этого у нас будет первый зарезервированный `memblock` для `EBDA` в секции `.meminit.data`. Функция `reserve_ebda_region` завершила работу над этим шагом, и мы можем вернуться в arch/x86/kernel/head64.c.

Мы закончили все приготовления! Последним шагом в функции `x86_64_start_reservations` является вызов функции `start_kernel` :

```
start_kernel()
```

расположенной в init/main.c.

Заключение

Это конец третей части инициализации ядра Linux. В следующей части мы увидим первые шаги инициализации в точке входа в ядро - `start_kernel` . Это будет первый шаг, прежде чем мы увидим запуск первого процесса `init` .

От переводчика: пожалуйста, имейте в виду, что английский - не мой родной язык, и я очень извиняюсь за возможные неудобства. Если вы найдёте какие-либо ошибки или неточности в переводе, пожалуйста, пришлите pull request в linux-insides-ru.

Ссылки

- BIOS data area
- Что такое Extended BIOS Data Area
- Предыдущая часть

Kernel initialization. Part 4.

Kernel entry point

If you have read the previous part - Last preparations before the kernel entry point, you can remember that we finished all pre-initialization stuff and stopped right before the call to the `start_kernel` function from the init/main.c. The `start_kernel` is the entry of the generic and architecture independent kernel code, although we will return to the `arch/` folder many times. If you look inside of the `start_kernel` function, you will see that this function is very big. For this moment it contains about `86` calls of functions. Yes, it's very big and of course this part will not cover all the processes that occur in this function. In the current part we will only start to do it. This part and all the next which will be in the Kernel initialization process chapter will cover it.

The main purpose of the `start_kernel` to finish kernel initialization process and launch the first `init` process. Before the first process will be started, the `start_kernel` must do many things such as: to enable lock validator, to initialize processor id, to enable early cgroups subsystem, to setup per-cpu areas, to initialize different caches in vfs, to initialize memory manager, rcu, vmalloc, scheduler, IRQs, ACPI and many many more. Only after these steps will we see the launch of the first `init` process in the last part of this chapter. So much kernel code awaits us, let's start.

NOTE: All parts from this big chapter `Linux Kernel initialization process` will not cover anything about debugging. There will be a separate chapter about kernel debugging tips.

A little about function attributes

As I wrote above, the `start_kernel` function is defined in the init/main.c. This function defined with the `__init` attribute and as you already may know from other parts, all functions which are defined with this attribute are necessary during kernel initialization.

```
#define __init      __section(.init.text) __cold notrace
```

After the initialization process have finished, the kernel will release these sections with a call to the `free_initmem` function. Note also that `__init` is defined with two attributes: `__cold` and `notrace`. The purpose of the first `cold` attribute is to mark that the function is rarely used and the compiler must optimize this function for size. The second `notrace` is defined as:

```
#define notrace __attribute__((no_instrument_function))
```

where `no_instrument_function` says to the compiler not to generate profiling function calls.

In the definition of the `start_kernel` function, you can also see the `__visible` attribute which expands to the:

```
#define __visible __attribute__((externally_visible))
```

where `externally_visible` tells to the compiler that something uses this function or variable, to prevent marking this function/variable as `unusable`. You can find the definition of this and other macro attributes in include/linux/init.h.

First steps in the start_kernel

At the beginning of the `start_kernel` you can see the definition of these two variables:

```
char *command_line;
```

```
char *after_dashes;
```

The first represents a pointer to the kernel command line and the second will contain the result of the `parse_args` function which parses an input string with parameters in the form `name=value`, looking for specific keywords and invoking the right handlers. We will not go into the details related with these two variables at this time, but will see it in the next parts. In the next step we can see a call to the `set_task_stack_end_magic` function. This function takes address of the `init_task` and sets `STACK_END_MAGIC` (`0x57AC6E9D`) as canary for it. `init_task` represents the initial task structure:

```
struct task_struct init_task = INIT_TASK(init_task);
```

where `task_struct` stores all the information about a process. I will not explain this structure in this book because it's very big. You can find its definition in include/linux/sched.h. At this moment `task_struct` contains more than `100` fields! Although you will not see the explanation of the `task_struct` in this book, we will use it very often since it is the fundamental structure which describes the `process` in the Linux kernel. I will describe the meaning of the fields of this structure as we meet them in practice.

You can see the definition of the `init_task` and it initialized by the `INIT_TASK` macro. This macro is from include/linux/init_task.h and it just fills the `init_task` with the values for the first process. For example it sets:

- init process state to zero or `runnable`. A runnable process is one which is waiting only for a CPU to run on;
- init process flags - `PF_KTHREAD` which means - kernel thread;
- a list of runnable task;
- process address space;
- init process stack to the `&init_thread_info` which is `init_thread_union.thread_info` and `initthread_union` has type - `thread_union` which contains `thread_info` and process stack:

```
union thread_union {
    struct thread_info thread_info;
    unsigned long stack[THREAD_SIZE/sizeof(long)];
};
```

Every process has its own stack and it is 16 kilobytes or 4 page frames. in `x86_64`. We can note that it is defined as array of `unsigned long`. The next field of the `thread_union` is - `thread_info` defined as:

```
struct thread_info {
        struct task_struct      *task;
        struct exec_domain      *exec_domain;
        __u32                   flags;
        __u32                   status;
        __u32                   cpu;
        int                     saved_preempt_count;
        mm_segment_t            addr_limit;
        struct restart_block    restart_block;
        void __user             *sysenter_return;
        unsigned int            sig_on_uaccess_error:1;
        unsigned int            uaccess_err:1;
};
```

and occupies 52 bytes. The `thread_info` structure contains architecture-specific information on the thread. We know that on `x86_64` the stack grows down and `thread_union.thread_info` is stored at the bottom of the stack in our case. So the process stack is 16 kilobytes and `thread_info` is at the bottom. The remaining thread_size will be `16 kilobytes - 62 bytes = 16332 bytes`. Note that `thread_union` represented as the union and not structure, it means that `thread_info` and stack share the memory space.

Schematically it can be represented as follows:

```
+-----------------------+
```

```
|                      |
|                      |
|        stack         |
|                      |
|______________________|
|          |           |
|          |           |
|          |           |
|__________↓___________|             +--------------------+
|                      |             |                    |
|      thread_info     |<----------->|     task_struct    |
|                      |             |                    |
+----------------------+             +--------------------+
```

http://www.quora.com/In-Linux-kernel-Why-thread_info-structure-and-the-kernel-stack-of-a-process-binds-in-union-construct

So the `INIT_TASK` macro fills these `task_struct's` fields and many many more. As I already wrote above, I will not describe all the fields and values in the `INIT_TASK` macro but we will see them soon.

Now let's go back to the `set_task_stack_end_magic` function. This function defined in the kernel/fork.c and sets a canary to the `init` process stack to prevent stack overflow.

```
void set_task_stack_end_magic(struct task_struct *tsk)
{
    unsigned long *stackend;
    stackend = end_of_stack(tsk);
    *stackend = STACK_END_MAGIC; /* for overflow detection */
}
```

Its implementation is simple. `set_task_stack_end_magic` gets the end of the stack for the given `task_struct` with the `end_of_stack` function. Earlier (and now for all architectures besides `x86_64`) stack was located in the `thread_info` structure. So the end of a process stack depends on the `CONFIG_STACK_GROWSUP` configuration option. As we learn in `x86_64` architecture, the stack grows down. So the end of the process stack will be:

```
(unsigned long *)(task_thread_info(p) + 1);
```

where `task_thread_info` just returns the stack which we filled with the `INIT_TASK` macro:

```
#define task_thread_info(task)  ((struct thread_info *)(task)->stack)
```

From the Linux kernel `v4.9-rc1` release, `thread_info` structure may contains only flags and stack pointer resides in `task_struct` structure which represents a thread in the Linux kernel. This depends on `CONFIG_THREAD_INFO_IN_TASK` kernel configuration option which is enabled by default for `x86_64` . You can be sure in this if you will look in the init/main.c configuration build file:

```
config THREAD_INFO_IN_TASK
    bool
    help
      Select this to move thread_info off the stack into task_struct.  To
      make this work, an arch will need to remove all thread_info fields
      except flags and fix any runtime bugs.

      One subtle change that will be needed is to use try_get_task_stack()
      and put_task_stack() in save_thread_stack_tsk() and get_wchan().
```

and arch/x86/Kconfig:

```
config X86
```

```
    def_bool y
        ...
        ...
        ...
        select THREAD_INFO_IN_TASK
        ...
        ...
        ...
```

So, in this way we may just get end of a thread stack from the given `task_struct` structure:

```
#ifdef CONFIG_THREAD_INFO_IN_TASK
static inline unsigned long *end_of_stack(const struct task_struct *task)
{
    return task->stack;
}
#endif
```

As we got the end of the init process stack, we write `STACK_END_MAGIC` there. After `canary` is set, we can check it like this:

```
if (*end_of_stack(task) != STACK_END_MAGIC) {
        //
        // handle stack overflow here
    //
}
```

The next function after the `set_task_stack_end_magic` is `smp_setup_processor_id`. This function has an empty body for `x86_64`:

```
void __init __weak smp_setup_processor_id(void)
{
}
```

as it not implemented for all architectures, but some such as s390 and arm64.

The next function in `start_kernel` is `debug_objects_early_init`. Implementation of this function is almost the same as `lockdep_init`, but fills hashes for object debugging. As I wrote above, we will not see the explanation of this and other functions which are for debugging purposes in this chapter.

After the `debug_object_early_init` function we can see the call of the `boot_init_stack_canary` function which fills `task_struct->canary` with the canary value for the `-fstack-protector` gcc feature. This function depends on the `CONFIG_CC_STACKPROTECTOR` configuration option and if this option is disabled, `boot_init_stack_canary` does nothing, otherwise it generates random numbers based on random pool and the TSC:

```
get_random_bytes(&canary, sizeof(canary));
tsc = __native_read_tsc();
canary += tsc + (tsc << 32UL);
```

After we got a random number, we fill the `stack_canary` field of `task_struct` with it:

```
current->stack_canary = canary;
```

and write this value to the top of the IRQ stack with the:

```
this_cpu_write(irq_stack_union.stack_canary, canary); // read below about this_cpu_write
```

Again, we will not dive into details here, we will cover it in the part about IRQs. As canary is set, we disable local and early boot IRQs and register the bootstrap CPU in the CPU maps. We disable local IRQs (interrupts for current CPU) with the `local_irq_disable` macro which expands to the call of the `arch_local_irq_disable` function from include/linux/percpu-defs.h:

```
static inline notrace void arch_local_irq_disable(void)
{
        native_irq_disable();
}
```

Where `native_irq_disable` is `cli` instruction for `x86_64`. As interrupts are disabled we can register the current CPU with the given ID in the CPU bitmap.

The first processor activation

The current function from the `start_kernel` is `boot_cpu_init`. This function initializes various CPU masks for the bootstrap processor. First of all it gets the bootstrap processor id with a call to:

```
int cpu = smp_processor_id();
```

For now it is just zero. If the `CONFIG_DEBUG_PREEMPT` configuration option is disabled, `smp_processor_id` just expands to the call of `raw_smp_processor_id` which expands to the:

```
#define raw_smp_processor_id() (this_cpu_read(cpu_number))
```

`this_cpu_read` as many other function like this (`this_cpu_write` , `this_cpu_add` and etc...) defined in the include/linux/percpu-defs.h and presents `this_cpu` operation. These operations provide a way of optimizing access to the per-cpu variables which are associated with the current processor. In our case it is `this_cpu_read` :

```
__pcpu_size_call_return(this_cpu_read_, pcp)
```

Remember that we have passed `cpu_number` as `pcp` to the `this_cpu_read` from the `raw_smp_processor_id` . Now let's look at the `__pcpu_size_call_return` implementation:

```
#define __pcpu_size_call_return(stem, variable)                         \
({                                                                      \
        typeof(variable) pscr_ret__;                                    \
        __verify_pcpu_ptr(&(variable));                                 \
        switch(sizeof(variable)) {                                      \
        case 1: pscr_ret__ = stem##1(variable); break;                  \
        case 2: pscr_ret__ = stem##2(variable); break;                  \
        case 4: pscr_ret__ = stem##4(variable); break;                  \
        case 8: pscr_ret__ = stem##8(variable); break;                  \
        default:                                                        \
                __bad_size_call_parameter(); break;                     \
        }                                                               \
        pscr_ret__;                                                     \
})
```

Yes, it looks a little strange but it's easy. First of all we can see the definition of the `pscr_ret__` variable with the `int` type. Why int? Ok, `variable` is `common_cpu` and it was declared as per-cpu int variable:

```
DECLARE_PER_CPU_READ_MOSTLY(int, cpu_number);
```

In the next step we call `__verify_pcpu_ptr` with the address of `cpu_number` . `__veryf_pcpu_ptr` used to verify that the given parameter is a per-cpu pointer. After that we set `pscr_ret__` value which depends on the size of the variable. Our `common_cpu` variable is `int` , so it 4 bytes in size. It means that we will get `this_cpu_read_4(common_cpu)` in `pscr_ret__` . In the end of the `__pcpu_size_call_return` we just call it. `this_cpu_read_4` is a macro:

```
#define this_cpu_read_4(pcp)       percpu_from_op("mov", pcp)
```

which calls `percpu_from_op` and pass `mov` instruction and per-cpu variable there. `percpu_from_op` will expand to the inline assembly call:

```
asm("movl %%gs:%1,%0" : "=r" (pfo_ret__) : "m" (common_cpu))
```

Let's try to understand how it works and what it does. The `gs` segment register contains the base of per-cpu area. Here we just copy `common_cpu` which is in memory to the `pfo_ret__` with the `movl` instruction. Or with another words:

```
this_cpu_read(common_cpu)
```

is the same as:

```
movl %gs:$common_cpu, $pfo_ret__
```

As we didn't setup per-cpu area, we have only one - for the current running CPU, we will get `zero` as a result of the `smp_processor_id` .

As we got the current processor id, `boot_cpu_init` sets the given CPU online, active, present and possible with the:

```
set_cpu_online(cpu, true);
set_cpu_active(cpu, true);
set_cpu_present(cpu, true);
set_cpu_possible(cpu, true);
```

All of these functions use the concept - `cpumask` . `cpu_possible` is a set of CPU ID's which can be plugged in at any time during the life of that system boot. `cpu_present` represents which CPUs are currently plugged in. `cpu_online` represents subset of the `cpu_present` and indicates CPUs which are available for scheduling. These masks depend on the `CONFIG_HOTPLUG_CPU` configuration option and if this option is disabled `possible == present` and `active == online` . Implementation of the all of these functions are very similar. Every function checks the second parameter. If it is `true` , it calls `cpumask_set_cpu` or `cpumask_clear_cpu` otherwise.

For example let's look at `set_cpu_possible` . As we passed `true` as the second parameter, the:

```
cpumask_set_cpu(cpu, to_cpumask(cpu_possible_bits));
```

will be called. First of all let's try to understand the `to_cpumask` macro. This macro casts a bitmap to a `struct cpumask *` . CPU masks provide a bitmap suitable for representing the set of CPU's in a system, one bit position per CPU number. CPU mask presented by the `cpu_mask` structure:

```
typedef struct cpumask { DECLARE_BITMAP(bits, NR_CPUS); } cpumask_t;
```

which is just bitmap declared with the `DECLARE_BITMAP` macro:

```
#define DECLARE_BITMAP(name, bits) unsigned long name[BITS_TO_LONGS(bits)]
```

As we can see from its definition, the `DECLARE_BITMAP` macro expands to the array of `unsigned long`. Now let's look at how the `to_cpumask` macro is implemented:

```
#define to_cpumask(bitmap)                                              \
        ((struct cpumask *)(1 ? (bitmap)                                \
                            : (void *)sizeof(__check_is_bitmap(bitmap))))
```

I don't know about you, but it looked really weird for me at the first time. We can see a ternary operator here which is `true` every time, but why the `__check_is_bitmap` here? It's simple, let's look at it:

```
static inline int __check_is_bitmap(const unsigned long *bitmap)
{
        return 1;
}
```

Yeah, it just returns `1` every time. Actually we need in it here only for one purpose: at compile time it checks that the given `bitmap` is a bitmap, or in other words it checks that the given `bitmap` has a type of `unsigned long *`. So we just pass `cpu_possible_bits` to the `to_cpumask` macro for converting the array of `unsigned long` to the `struct cpumask *`. Now we can call `cpumask_set_cpu` function with the `cpu` - 0 and `struct cpumask *cpu_possible_bits`. This function makes only one call of the `set_bit` function which sets the given `cpu` in the cpumask. All of these `set_cpu_*` functions work on the same principle.

If you're not sure that this `set_cpu_*` operations and `cpumask` are not clear for you, don't worry about it. You can get more info by reading the special part about it - cpumask or documentation.

As we activated the bootstrap processor, it's time to go to the next function in the `start_kernel.` Now it is `page_address_init`, but this function does nothing in our case, because it executes only when all `RAM` can't be mapped directly.

Print linux banner

The next call is `pr_notice`:

```
#define pr_notice(fmt, ...) \
    printk(KERN_NOTICE pr_fmt(fmt), ##__VA_ARGS__)
```

as you can see it just expands to the `printk` call. At this moment we use `pr_notice` to print the Linux banner:

```
pr_notice("%s", linux_banner);
```

which is just the kernel version with some additional parameters:

```
Linux version 4.0.0-rc6+ (alex@localhost) (gcc version 4.9.1 (Ubuntu 4.9.1-16ubuntu6) ) #319 SMP
```

Architecture-dependent parts of initialization

The next step is architecture-specific initialization. The Linux kernel does it with the call of the `setup_arch` function. This is a very big function like `start_kernel` and we do not have time to consider all of its implementation in this part. Here we'll only start to do it and continue in the next part. As it is `architecture-specific`, we need to go again to the `arch/` directory. The `setup_arch` function defined in the arch/x86/kernel/setup.c source code file and takes only one argument - address of the kernel command line.

This function starts from the reserving memory block for the kernel `_text` and `_data` which starts from the `_text` symbol (you can remember it from the arch/x86/kernel/head_64.S) and ends before `__bss_stop`. We are using `memblock` for the reserving of memory block:

```
memblock_reserve(__pa_symbol(_text), (unsigned long)__bss_stop - (unsigned long)_text);
```

You can read about `memblock` in the Linux kernel memory management Part 1.. As you can remember `memblock_reserve` function takes two parameters:

- base physical address of a memory block;
- size of a memory block.

We can get the base physical address of the `_text` symbol with the `__pa_symbol` macro:

```
#define __pa_symbol(x) \
    __phys_addr_symbol(__phys_reloc_hide((unsigned long)(x)))
```

First of all it calls `__phys_reloc_hide` macro on the given parameter. The `__phys_reloc_hide` macro does nothing for `x86_64` and just returns the given parameter. Implementation of the `__phys_addr_symbol` macro is easy. It just subtracts the symbol address from the base address of the kernel text mapping base virtual address (you can remember that it is `__START_KERNEL_map`) and adds `phys_base` which is the base address of `_text`:

```
#define __phys_addr_symbol(x) \
 ((unsigned long)(x) - __START_KERNEL_map + phys_base)
```

After we got the physical address of the `_text` symbol, `memblock_reserve` can reserve a memory block from the `_text` to the `__bss_stop - _text`.

Reserve memory for initrd

In the next step after we reserved place for the kernel text and data is reserving place for the initrd. We will not see details about `initrd` in this post, you just may know that it is temporary root file system stored in memory and used by the kernel during its startup. The `early_reserve_initrd` function does all work. First of all this function gets the base address of the ram disk, its size and the end address with:

```
u64 ramdisk_image = get_ramdisk_image();
u64 ramdisk_size  = get_ramdisk_size();
u64 ramdisk_end   = PAGE_ALIGN(ramdisk_image + ramdisk_size);
```

All of these parameters are taken from `boot_params`. If you have read the chapter about Linux Kernel Booting Process, you must remember that we filled the `boot_params` structure during boot time. The kernel setup header contains a couple of fields which describes ramdisk, for example:

```
Field name:    ramdisk_image
Type:        write (obligatory)
Offset/size:    0x218/4
Protocol:    2.00+

  The 32-bit linear address of the initial ramdisk or ramfs.  Leave at
  zero if there is no initial ramdisk/ramfs.
```

So we can get all the information that interests us from `boot_params`. For example let's look at `get_ramdisk_image`:

```
static u64 __init get_ramdisk_image(void)
{
        u64 ramdisk_image = boot_params.hdr.ramdisk_image;

        ramdisk_image |= (u64)boot_params.ext_ramdisk_image << 32;

        return ramdisk_image;
}
```

Here we get the address of the ramdisk from the `boot_params` and shift left it on `32`. We need to do it because as you can read in the Documentation/x86/zero-page.txt:

```
0C0/004    ALL    ext_ramdisk_image ramdisk_image high 32bits
```

So after shifting it on 32, we're getting a 64-bit address in `ramdisk_image` and we return it. `get_ramdisk_size` works on the same principle as `get_ramdisk_image`, but it used `ext_ramdisk_size` instead of `ext_ramdisk_image`. After we got ramdisk's size, base address and end address, we check that bootloader provided ramdisk with the:

```
if (!boot_params.hdr.type_of_loader ||
    !ramdisk_image || !ramdisk_size)
    return;
```

and reserve memory block with the calculated addresses for the initial ramdisk in the end:

```
memblock_reserve(ramdisk_image, ramdisk_end - ramdisk_image);
```

Conclusion

It is the end of the fourth part about the Linux kernel initialization process. We started to dive in the kernel generic code from the `start_kernel` function in this part and stopped on the architecture-specific initialization in the `setup_arch`. In the next part we will continue with architecture-dependent initialization steps.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me a PR to linux-insides.

Links

- GCC function attributes
- this_cpu operations
- cpumask
- lock validator
- cgroups
- stack buffer overflow
- IRQs
- initrd
- Previous part

Kernel initialization. Part 5.

Continue of architecture-specific initialization

In the previous part, we stopped at the initialization of an architecture-specific stuff from the setup_arch function and now we will continue with it. As we reserved memory for the initrd, next step is the `olpc_ofw_detect` which detects One Laptop Per Child support. We will not consider platform related stuff in this book and will skip functions related with it. So let's go ahead. The next step is the `early_trap_init` function. This function initializes debug (`#DB` - raised when the `TF` flag of rflags is set) and `int3` (`#BP`) interrupts gate. If you don't know anything about interrupts, you can read about it in the Early interrupt and exception handling. In `x86` architecture `INT` , `INTO` and `INT3` are special instructions which allow a task to explicitly call an interrupt handler. The `INT3` instruction calls the breakpoint (`#BP`) handler. You may remember, we already saw it in the part about interrupts: and exceptions:

```
----------------------------------------------------------------------------------------------
|Vector|Mnemonic|Description         |Type |Error Code|Source                   |
----------------------------------------------------------------------------------------------
|3     | #BP    |Breakpoint          |Trap |NO        |INT 3                    |
----------------------------------------------------------------------------------------------
```

Debug interrupt `#DB` is the primary method of invoking debuggers. `early_trap_init` defined in the arch/x86/kernel/traps.c. This functions sets `#DB` and `#BP` handlers and reloads IDT:

```
void __init early_trap_init(void)
{
        set_intr_gate_ist(X86_TRAP_DB, &debug, DEBUG_STACK);
        set_system_intr_gate_ist(X86_TRAP_BP, &int3, DEBUG_STACK);
        load_idt(&idt_descr);
}
```

We already saw implementation of the `set_intr_gate` in the previous part about interrupts. Here are two similar functions `set_intr_gate_ist` and `set_system_intr_gate_ist` . Both of these two functions take three parameters:

- number of the interrupt;
- base address of the interrupt/exception handler;
- third parameter is - `Interrupt Stack Table` . `IST` is a new mechanism in the `x86_64` and part of the TSS. Every active thread in kernel mode has own kernel stack which is `16` kilobytes. While a thread in user space, this kernel stack is empty.

In addition to per-thread stacks, there are a couple of specialized stacks associated with each CPU. All about these stack you can read in the linux kernel documentation - Kernel stacks. `x86_64` provides feature which allows to switch to a new `special` stack for during any events as non-maskable interrupt and etc... And the name of this feature is - `Interrupt Stack Table` . There can be up to 7 `IST` entries per CPU and every entry points to the dedicated stack. In our case this is `DEBUG_STACK` .

`set_intr_gate_ist` and `set_system_intr_gate_ist` work by the same principle as `set_intr_gate` with only one difference. Both of these functions checks interrupt number and call `_set_gate` inside:

```
BUG_ON((unsigned)n > 0xFF);
_set_gate(n, GATE_INTERRUPT, addr, 0, ist, __KERNEL_CS);
```

as `set_intr_gate` does this. But `set_intr_gate` calls `_set_gate` with dpl - 0, and ist - 0, but `set_intr_gate_ist` and `set_system_intr_gate_ist` sets `ist` as `DEBUG_STACK` and `set_system_intr_gate_ist` sets `dpl` as `0x3` which is the lowest privilege. When an interrupt occurs and the hardware loads such a descriptor, then hardware automatically sets the new stack pointer based on the IST value, then invokes the interrupt handler. All of the special kernel stacks will be set in the `cpu_init` function (we will see it later).

As `#DB` and `#BP` gates written to the `idt_descr`, we reload `IDT` table with `load_idt` which just calss `ldtr` instruction. Now let's look on interrupt handlers and will try to understand how they works. Of course, I can't cover all interrupt handlers in this book and I do not see the point in this. It is very interesting to delve in the linux kernel source code, so we will see how `debug` handler implemented in this part, and understand how other interrupt handlers are implemented will be your task.

DB handler

As you can read above, we passed address of the `#DB` handler as `&debug` in the `set_intr_gate_ist`. lxr.free-electrons.com is a great resource for searching identifiers in the linux kernel source code, but unfortunately you will not find `debug` handler with it. All of you can find, it is `debug` definition in the arch/x86/include/asm/traps.h:

```
asmlinkage void debug(void);
```

We can see `asmlinkage` attribute which tells to us that `debug` is function written with assembly. Yeah, again and again assembly :). Implementation of the `#DB` handler as other handlers is in this arch/x86/entry/entry_64.S and defined with the `idtentry` assembly macro:

```
idtentry debug do_debug has_error_code=0 paranoid=1 shift_ist=DEBUG_STACK
```

`idtentry` is a macro which defines an interrupt/exception entry point. As you can see it takes five arguments:

- name of the interrupt entry point;
- name of the interrupt handler;
- has interrupt error code or not;
- paranoid - if this parameter = 1, switch to special stack (read above);
- shift_ist - stack to switch during interrupt.

Now let's look on `idtentry` macro implementation. This macro defined in the same assembly file and defines `debug` function with the `ENTRY` macro. For the start `idtentry` macro checks that given parameters are correct in case if need to switch to the special stack. In the next step it checks that give interrupt returns error code. If interrupt does not return error code (in our case `#DB` does not return error code), it calls `INTR_FRAME` or `XCPT_FRAME` if interrupt has error code. Both of these macros `XCPT_FRAME` and `INTR_FRAME` do nothing and need only for the building initial frame state for interrupts. They uses `CFI` directives and used for debugging. More info you can find in the CFI directives. As comment from the arch/x86/kernel/entry_64.S says: `CFI macros are used to generate dwarf2 unwind information for better backtraces. They don't change any code.` so we will ignore them.

```
.macro idtentry sym do_sym has_error_code:req paranoid=0 shift_ist=-1
ENTRY(\sym)
    /* Sanity check */
    .if \shift_ist != -1 && \paranoid == 0
    .error "using shift_ist requires paranoid=1"
    .endif

    .if \has_error_code
    XCPT_FRAME
    .else
    INTR_FRAME
    .endif
    ...
    ...
    ...
```

You can remember from the previous part about early interrupts/exceptions handling that after interrupt occurs, current stack will have following format:

```
    +-----------------------+
    |                       |
+40 |         SS            |
+32 |         RSP           |
+24 |        RFLAGS         |
+16 |         CS            |
+8  |         RIP           |
 0  |       Error Code      | <---- rsp
    |                       |
    +-----------------------+
```

The next two macro from the `idtentry` implementation are:

```
    ASM_CLAC
    PARAVIRT_ADJUST_EXCEPTION_FRAME
```

First `ASM_CLAC` macro depends on `CONFIG_X86_SMAP` configuration option and need for security reason, more about it you can read here. The second `PARAVIRT_ADJUST_EXCEPTION_FRAME` macro is for handling handle Xen-type-exceptions (this chapter about kernel initialization and we will not consider virtualization stuff here).

The next piece of code checks if interrupt has error code or not and pushes `$-1` which is `0xffffffffffffffff` on `x86_64` on the stack if not:

```
    .ifeq \has_error_code
    pushq_cfi $-1
    .endif
```

We need to do it as `dummy` error code for stack consistency for all interrupts. In the next step we subtract from the stack pointer `$ORIG_RAX-R15` :

```
    subq $ORIG_RAX-R15, %rsp
```

where `ORIRG_RAX` , `R15` and other macros defined in the arch/x86/include/asm/calling.h and `ORIG_RAX-R15` is 120 bytes. General purpose registers will occupy these 120 bytes because we need to store all registers on the stack during interrupt handling. After we set stack for general purpose registers, the next step is checking that interrupt came from userspace with:

```
testl $3, CS(%rsp)
jnz 1f
```

Here we checks first and second bits in the `CS` . You can remember that `CS` register contains segment selector where first two bits are `RPL` . All privilege levels are integers in the range 0–3, where the lowest number corresponds to the highest privilege. So if interrupt came from the kernel mode we call `save_paranoid` or jump on label `1` if not. In the `save_paranoid` we store all general purpose registers on the stack and switch user `gs` on kernel `gs` if need:

```
    movl $1,%ebx
    movl $MSR_GS_BASE,%ecx
    rdmsr
    testl %edx,%edx
    js 1f
    SWAPGS
    xorl %ebx,%ebx
1:    ret
```

In the next steps we put `pt_regs` pointer to the `rdi`, save error code in the `rsi` if it has and call interrupt handler which is - `do_debug` in our case from the arch/x86/kernel/traps.c. `do_debug` like other handlers takes two parameters:

- pt_regs - is a structure which presents set of CPU registers which are saved in the process' memory region;
- error code - error code of interrupt.

After interrupt handler finished its work, calls `paranoid_exit` which restores stack, switch on userspace if interrupt came from there and calls `iret`. That's all. Of course it is not all :), but we will see more deeply in the separate chapter about interrupts.

This is general view of the `idtentry` macro for `#DB` interrupt. All interrupts are similar to this implementation and defined with idtentry too. After `early_trap_init` finished its work, the next function is `early_cpu_init`. This function defined in the arch/x86/kernel/cpu/common.c and collects information about CPU and its vendor.

Early ioremap initialization

The next step is initialization of early `ioremap`. In general there are two ways to communicate with devices:

- I/O Ports;
- Device memory.

We already saw first method (`outb/inb` instructions) in the part about linux kernel booting process. The second method is to map I/O physical addresses to virtual addresses. When a physical address is accessed by the CPU, it may refer to a portion of physical RAM which can be mapped on memory of the I/O device. So `ioremap` used to map device memory into kernel address space.

As i wrote above next function is the `early_ioremap_init` which re-maps I/O memory to kernel address space so it can access it. We need to initialize early ioremap for early initialization code which needs to temporarily map I/O or memory regions before the normal mapping functions like `ioremap` are available. Implementation of this function is in the arch/x86/mm/ioremap.c. At the start of the `early_ioremap_init` we can see definition of the `pmd` point with `pmd_t` type (which presents page middle directory entry `typedef struct { pmdval_t pmd; } pmd_t;` where `pmdval_t` is `unsigned long`) and make a check that `fixmap` aligned in a correct way:

```
pmd_t *pmd;
BUILD_BUG_ON((fix_to_virt(0) + PAGE_SIZE) & ((1 << PMD_SHIFT) - 1));
```

`fixmap` - is fixed virtual address mappings which extends from `FIXADDR_START` to `FIXADDR_TOP`. Fixed virtual addresses are needed for subsystems that need to know the virtual address at compile time. After the check `early_ioremap_init` makes a call of the `early_ioremap_setup` function from the mm/early_ioremap.c. `early_ioremap_setup` fills `slot_virt` array of the `unsigned long` with virtual addresses with 512 temporary boot-time fix-mappings:

```
for (i = 0; i < FIX_BTMAPS_SLOTS; i++)
    slot_virt[i] = __fix_to_virt(FIX_BTMAP_BEGIN - NR_FIX_BTMAPS*i);
```

After this we get page middle directory entry for the `FIX_BTMAP_BEGIN` and put to the `pmd` variable, fills `bm_pte` with zeros which is boot time page tables and call `pmd_populate_kernel` function for setting given page table entry in the given page middle directory:

```
pmd = early_ioremap_pmd(fix_to_virt(FIX_BTMAP_BEGIN));
memset(bm_pte, 0, sizeof(bm_pte));
pmd_populate_kernel(&init_mm, pmd, bm_pte);
```

That's all for this. If you feeling puzzled, don't worry. There is special part about `ioremap` and `fixmaps` in the Linux Kernel Memory Management. Part 2 chapter.

Obtaining major and minor numbers for the root device

After early `ioremap` was initialized, you can see the following code:

```
ROOT_DEV = old_decode_dev(boot_params.hdr.root_dev);
```

This code obtains major and minor numbers for the root device where `initrd` will be mounted later in the `do_mount_root` function. Major number of the device identifies a driver associated with the device. Minor number referred on the device controlled by driver. Note that `old_decode_dev` takes one parameter from the `boot_params_structure`. As we can read from the x86 linux kernel boot protocol:

```
Field name:    root_dev
Type:          modify (optional)
Offset/size:    0x1fc/2
Protocol:       ALL

  The default root device device number.  The use of this field is
  deprecated, use the "root=" option on the command line instead.
```

Now let's try to understand what `old_decode_dev` does. Actually it just calls `MKDEV` inside which generates `dev_t` from the give major and minor numbers. It's implementation is pretty simple:

```
static inline dev_t old_decode_dev(u16 val)
{
         return MKDEV((val >> 8) & 255, val & 255);
}
```

where `dev_t` is a kernel data type to present major/minor number pair. But what's the strange `old_` prefix? For historical reasons, there are two ways of managing the major and minor numbers of a device. In the first way major and minor numbers occupied 2 bytes. You can see it in the previous code: 8 bit for major number and 8 bit for minor number. But there is a problem: only 256 major numbers and 256 minor numbers are possible. So 16-bit integer was replaced by 32-bit integer where 12 bits reserved for major number and 20 bits for minor. You can see this in the `new_decode_dev` implementation:

```
static inline dev_t new_decode_dev(u32 dev)
{
         unsigned major = (dev & 0xfff00) >> 8;
         unsigned minor = (dev & 0xff) | ((dev >> 12) & 0xfff00);
         return MKDEV(major, minor);
}
```

After calculation we will get `0xfff` or 12 bits for `major` if it is `0xffffffff` and `0xfffff` or 20 bits for `minor`. So in the end of execution of the `old_decode_dev` we will get major and minor numbers for the root device in `ROOT_DEV`.

Memory map setup

The next point is the setup of the memory map with the call of the `setup_memory_map` function. But before this we setup different parameters as information about a screen (current row and column, video page and etc... (you can read about it in the Video mode initialization and transition to protected mode)), Extended display identification data, video mode, bootloader_type and etc...:

```
    screen_info = boot_params.screen_info;
    edid_info = boot_params.edid_info;
    saved_video_mode = boot_params.hdr.vid_mode;
    bootloader_type = boot_params.hdr.type_of_loader;
    if ((bootloader_type >> 4) == 0xe) {
        bootloader_type &= 0xf;
```

```
        bootloader_type |= (boot_params.hdr.ext_loader_type+0x10) << 4;
    }
    bootloader_version  = bootloader_type & 0xf;
    bootloader_version |= boot_params.hdr.ext_loader_ver << 4;
```

All of these parameters we got during boot time and stored in the `boot_params` structure. After this we need to setup the end of the I/O memory. As you know one of the main purposes of the kernel is resource management. And one of the resource is memory. As we already know there are two ways to communicate with devices are I/O ports and device memory. All information about registered resources are available through:

- /proc/ioports - provides a list of currently registered port regions used for input or output communication with a device;
- /proc/iomem - provides current map of the system's memory for each physical device.

At the moment we are interested in `/proc/iomem` :

```
cat /proc/iomem
00000000-00000fff : reserved
00001000-0009d7ff : System RAM
0009d800-0009ffff : reserved
000a0000-000bffff : PCI Bus 0000:00
000c0000-000cffff : Video ROM
000d0000-000d3fff : PCI Bus 0000:00
000d4000-000d7fff : PCI Bus 0000:00
000d8000-000dbfff : PCI Bus 0000:00
000dc000-000dffff : PCI Bus 0000:00
000e0000-000fffff : reserved
  000e0000-000e3fff : PCI Bus 0000:00
  000e4000-000e7fff : PCI Bus 0000:00
  000f0000-000fffff : System ROM
```

As you can see range of addresses are shown in hexadecimal notation with its owner. Linux kernel provides API for managing any resources in a general way. Global resources (for example PICs or I/O ports) can be divided into subsets - relating to any hardware bus slot. The main structure `resource` :

```
struct resource {
        resource_size_t start;
        resource_size_t end;
        const char *name;
        unsigned long flags;
        struct resource *parent, *sibling, *child;
};
```

presents abstraction for a tree-like subset of system resources. This structure provides range of addresses from `start` to `end` (`resource_size_t` is `phys_addr_t` or `u64` for `x86_64`) which a resource covers, `name` of a resource (you see these names in the `/proc/iomem` output) and `flags` of a resource (All resources flags defined in the include/linux/ioport.h). The last are three pointers to the `resource` structure. These pointers enable a tree-like structure:

```
+-------------+      +-------------+
|             |      |             |
|    parent   |------|    sibling  |
|             |      |             |
+-------------+      +-------------+
       |
       |
+-------------+
|             |
|    child    |
|             |
+-------------+
```

Every subset of resources has root range resources. For `iomem` it is `iomem_resource` which defined as:

```
struct resource iomem_resource = {
        .name   = "PCI mem",
        .start  = 0,
        .end    = -1,
        .flags  = IORESOURCE_MEM,
};
EXPORT_SYMBOL(iomem_resource);
```

TODO EXPORT_SYMBOL

`iomem_resource` defines root addresses range for io memory with `PCI mem` name and `IORESOURCE_MEM` (`0x00000200`) as flags. As i wrote above our current point is setup the end address of the `iomem` . We will do it with:

```
iomem_resource.end = (1ULL << boot_cpu_data.x86_phys_bits) - 1;
```

Here we shift `1` on `boot_cpu_data.x86_phys_bits` . `boot_cpu_data` is `cpuinfo_x86` structure which we filled during execution of the `early_cpu_init` . As you can understand from the name of the `x86_phys_bits` field, it presents maximum bits amount of the maximum physical address in the system. Note also that `iomem_resource` is passed to the `EXPORT_SYMBOL` macro. This macro exports the given symbol (`iomem_resource` in our case) for dynamic linking or in other words it makes a symbol accessible to dynamically loaded modules.

After we set the end address of the root `iomem` resource address range, as I wrote above the next step will be setup of the memory map. It will be produced with the call of the `setup_ memory_map` function:

```
void __init setup_memory_map(void)
{
        char *who;

        who = x86_init.resources.memory_setup();
        memcpy(&e820_saved, &e820, sizeof(struct e820map));
        printk(KERN_INFO "e820: BIOS-provided physical RAM map:\n");
        e820_print_map(who);
}
```

First of all we call look here the call of the `x86_init.resources.memory_setup` . `x86_init` is a `x86_init_ops` structure which presents platform specific setup functions as resources initialization, pci initialization and etc... initialization of the `x86_init` is in the arch/x86/kernel/x86_init.c. I will not give here the full description because it is very long, but only one part which interests us for now:

```
struct x86_init_ops x86_init __initdata = {
    .resources = {
            .probe_roms             = probe_roms,
            .reserve_resources      = reserve_standard_io_resources,
            .memory_setup           = default_machine_specific_memory_setup,
    },
    ...
    ...
    ...
}
```

As we can see here `memry_setup` field is `default_machine_specific_memory_setup` where we get the number of the e820 entries which we collected in the boot time, sanitize the BIOS e820 map and fill `e820map` structure with the memory regions. As all regions are collected, print of all regions with printk. You can find this print if you execute `dmesg` command and you can see something like this:

```
[    0.000000] e820: BIOS-provided physical RAM map:
```

```
[    0.000000] BIOS-e820: [mem 0x0000000000000000-0x000000000009d7ff] usable
[    0.000000] BIOS-e820: [mem 0x000000000009d800-0x000000000009ffff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000000e0000-0x00000000000fffff] reserved
[    0.000000] BIOS-e820: [mem 0x0000000000100000-0x00000000be825fff] usable
[    0.000000] BIOS-e820: [mem 0x00000000be826000-0x00000000be82cfff] ACPI NVS
[    0.000000] BIOS-e820: [mem 0x00000000be82d000-0x00000000bf744fff] usable
[    0.000000] BIOS-e820: [mem 0x00000000bf745000-0x00000000bfff4fff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000bfff5000-0x00000000dc041fff] usable
[    0.000000] BIOS-e820: [mem 0x00000000dc042000-0x00000000dc0d2fff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000dc0d3000-0x00000000dc138fff] usable
[    0.000000] BIOS-e820: [mem 0x00000000dc139000-0x00000000dc27dfff] ACPI NVS
[    0.000000] BIOS-e820: [mem 0x00000000dc27e000-0x00000000deffefff] reserved
[    0.000000] BIOS-e820: [mem 0x00000000defff000-0x00000000deffffff] usable
...
...
...
```

Copying of the BIOS Enhanced Disk Device information

The next two steps is parsing of the `setup_data` with `parse_setup_data` function and copying BIOS EDD to the safe place. `setup_data` is a field from the kernel boot header and as we can read from the `x86` boot protocol:

```
Field name:    setup_data
Type:        write (special)
Offset/size:    0x250/8
Protocol:    2.09+

  The 64-bit physical pointer to NULL terminated single linked list of
  struct setup_data. This is used to define a more extensible boot
  parameters passing mechanism.
```

It used for storing setup information for different types as device tree blob, EFI setup data and etc... In the second step we copy BIOS EDD information from the `boot_params` structure that we collected in the arch/x86/boot/edd.c to the `edd` structure:

```
static inline void __init copy_edd(void)
{
     memcpy(edd.mbr_signature, boot_params.edd_mbr_sig_buffer,
            sizeof(edd.mbr_signature));
     memcpy(edd.edd_info, boot_params.eddbuf, sizeof(edd.edd_info));
     edd.mbr_signature_nr = boot_params.edd_mbr_sig_buf_entries;
     edd.edd_info_nr = boot_params.eddbuf_entries;
}
```

Memory descriptor initialization

The next step is initialization of the memory descriptor of the init process. As you already can know every process has its own address space. This address space presented with special data structure which called `memory descriptor`. Directly in the linux kernel source code memory descriptor presented with `mm_struct` structure. `mm_struct` contains many different fields related with the process address space as start/end address of the kernel code/data, start/end of the brk, number of memory areas, list of memory areas and etc... This structure defined in the include/linux/mm_types.h. As every process has its own memory descriptor, `task_struct` structure contains it in the `mm` and `active_mm` field. And our first `init` process has it too. You can remember that we saw the part of initialization of the init `task_struct` with `INIT_TASK` macro in the previous part:

```
#define INIT_TASK(tsk)  \
{
    ...
    ...
```

```
    ...
    .mm = NULL,          \
    .active_mm  = &init_mm, \
    ...
}
```

`mm` points to the process address space and `active_mm` points to the active address space if process has no address space such as kernel threads (more about it you can read in the documentation). Now we fill memory descriptor of the initial process:

```
    init_mm.start_code = (unsigned long) _text;
    init_mm.end_code = (unsigned long) _etext;
    init_mm.end_data = (unsigned long) _edata;
    init_mm.brk = _brk_end;
```

with the kernel's text, data and brk. `init_mm` is the memory descriptor of the initial process and defined as:

```
struct mm_struct init_mm = {
    .mm_rb          = RB_ROOT,
    .pgd            = swapper_pg_dir,
    .mm_users       = ATOMIC_INIT(2),
    .mm_count       = ATOMIC_INIT(1),
    .mmap_sem       = __RWSEM_INITIALIZER(init_mm.mmap_sem),
    .page_table_lock =  __SPIN_LOCK_UNLOCKED(init_mm.page_table_lock),
    .mmlist         = LIST_HEAD_INIT(init_mm.mmlist),
    INIT_MM_CONTEXT(init_mm)
};
```

where `mm_rb` is a red-black tree of the virtual memory areas, `pgd` is a pointer to the page global directory, `mm_users` is address space users, `mm_count` is primary usage counter and `mmap_sem` is memory area semaphore. After we setup memory descriptor of the initial process, next step is initialization of the Intel Memory Protection Extensions with `mpx_mm_init`. The next step is initialization of the code/data/bss resources with:

```
    code_resource.start = __pa_symbol(_text);
    code_resource.end = __pa_symbol(_etext)-1;
    data_resource.start = __pa_symbol(_etext);
    data_resource.end = __pa_symbol(_edata)-1;
    bss_resource.start = __pa_symbol(__bss_start);
    bss_resource.end = __pa_symbol(__bss_stop)-1;
```

We already know a little about `resource` structure (read above). Here we fills code/data/bss resources with their physical addresses. You can see it in the `/proc/iomem` :

```
00100000-be825fff : System RAM
  01000000-015bb392 : Kernel code
  015bb393-01930c3f : Kernel data
  01a11000-01ac3fff : Kernel bss
```

All of these structures are defined in the arch/x86/kernel/setup.c and look like typical resource initialization:

```
static struct resource code_resource = {
    .name    = "Kernel code",
    .start    = 0,
    .end    = 0,
    .flags    = IORESOURCE_BUSY | IORESOURCE_MEM
};
```

The last step which we will cover in this part will be `NX` configuration. `NX-bit` or no execute bit is 63-bit in the page directory entry which controls the ability to execute code from all physical pages mapped by the table entry. This bit can only be used/set when the `no-execute` page-protection mechanism is enabled by the setting `EFER.NXE` to 1. In the `x86_configure_nx` function we check that CPU has support of `NX-bit` and it does not disabled. After the check we fill `__supported_pte_mask` depend on it:

```
void x86_configure_nx(void)
{
        if (cpu_has_nx && !disable_nx)
                __supported_pte_mask |= _PAGE_NX;
        else
                __supported_pte_mask &= ~_PAGE_NX;
}
```

Conclusion

It is the end of the fifth part about linux kernel initialization process. In this part we continued to dive in the `setup_arch` function which makes initialization of architecture-specific stuff. It was long part, but we have not finished with it. As i already wrote, the `setup_arch` is big function, and I am really not sure that we will cover all of it even in the next part. There were some new interesting concepts in this part like `Fix-mapped` addresses, ioremap and etc... Don't worry if they are unclear for you. There is a special part about these concepts - Linux kernel memory management Part 2.. In the next part we will continue with the initialization of the architecture-specific stuff and will see parsing of the early kernel parameters, early dump of the pci devices, `Desktop Management Interface` scanning and many many more.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- mm vs active_mm
- e820
- Supervisor mode access prevention
- Kernel stacks
- TSS
- IDT
- Memory mapped I/O
- CFI directives
- PDF. dwarf4 specification
- Call stack
- Previous part

Kernel initialization. Part 6.

Architecture-specific initialization, again...

In the previous part we saw architecture-specific (`x86_64` in our case) initialization stuff from the arch/x86/kernel/setup.c and finished on `x86_configure_nx` function which sets the `_PAGE_NX` flag depends on support of NX bit. As I wrote before `setup_arch` function and `start_kernel` are very big, so in this and in the next part we will continue to learn about architecture-specific initialization process. The next function after `x86_configure_nx` is `parse_early_param` . This function is defined in the init/main.c and as you can understand from its name, this function parses kernel command line and setups different services depends on the given parameters (all kernel command line parameters you can find are in the Documentation/kernel-parameters.txt). You may remember how we setup `earlyprintk` in the earliest part. On the early stage we looked for kernel parameters and their value with the `cmdline_find_option` function and `__cmdline_find_option` , `__cmdline_find_option_bool` helpers from the arch/x86/boot/cmdline.c. There we're in the generic kernel part which does not depend on architecture and here we use another approach. If you are reading linux kernel source code, you already note calls like this:

```
early_param("gbpages", parse_direct_gbpages_on);
```

`early_param` macro takes two parameters:

- command line parameter name;
- function which will be called if given parameter is passed.

and defined as:

```
#define early_param(str, fn) \
        __setup_param(str, fn, fn, 1)
```

in the include/linux/init.h. As you can see `early_param` macro just makes call of the `__setup_param` macro:

```
#define __setup_param(str, unique_id, fn, early)                \
        static const char __setup_str_##unique_id[] __initconst \
                __aligned(1) = str; \
        static struct obs_kernel_param __setup_##unique_id      \
                __used __section(.init.setup)                   \
                __attribute__((aligned((sizeof(long)))))        \
                = { __setup_str_##unique_id, fn, early }
```

This macro defines `__setup_str_*_id` variable (where `*` depends on given function name) and assigns it to the given command line parameter name. In the next line we can see definition of the `__setup_*` variable which type is `obs_kernel_param` and its initialization. `obs_kernel_param` structure defined as:

```
struct obs_kernel_param {
        const char *str;
        int (*setup_func)(char *);
        int early;
};
```

and contains three fields:

- name of the kernel parameter;
- function which setups something depend on parameter;

* field determines is parameter early (1) or not (0).

Note that `__set_param` macro defines with `__section(.init.setup)` attribute. It means that all `__setup_str_*` will be placed in the `.init.setup` section, moreover, as we can see in the include/asm-generic/vmlinux.lds.h, they will be placed between `__setup_start` and `__setup_end` :

```
#define INIT_SETUP(initsetup_align)                \
                . = ALIGN(initsetup_align);        \
                VMLINUX_SYMBOL(__setup_start) = .; \
                *(.init.setup)                     \
                VMLINUX_SYMBOL(__setup_end) = .;
```

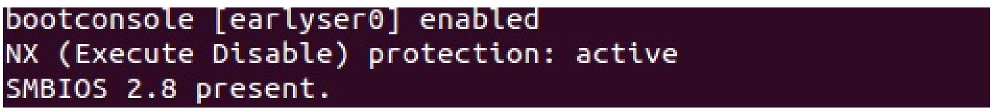
Now we know how parameters are defined, let's back to the `parse_early_param` implementation:

```
void __init parse_early_param(void)
{
        static int done __initdata;
        static char tmp_cmdline[COMMAND_LINE_SIZE] __initdata;

        if (done)
                return;

        /* All fall through to do_early_param. */
        strlcpy(tmp_cmdline, boot_command_line, COMMAND_LINE_SIZE);
        parse_early_options(tmp_cmdline);
        done = 1;
}
```

The `parse_early_param` function defines two static variables. First `done` check that `parse_early_param` already called and the second is temporary storage for kernel command line. After this we copy `boot_command_line` to the temporary command line which we just defined and call the `parse_early_options` function from the same source code `main.c` file. `parse_early_options` calls the `parse_args` function from the kernel/params.c where `parse_args` parses given command line and calls `do_early_param` function. This function goes from the `__setup_start` to `__setup_end` , and calls the function from the `obs_kernel_param` if a parameter is early. After this all services which are depend on early command line parameters were setup and the next call after the `parse_early_param` is `x86_report_nx` . As I wrote in the beginning of this part, we already set `NX-bit` with the `x86_configure_nx` . The next `x86_report_nx` function from the arch/x86/mm/setup_nx.c just prints information about the `NX` . Note that we call `x86_report_nx` not right after the `x86_configure_nx` , but after the call of the `parse_early_param` . The answer is simple: we call it after the `parse_early_param` because the kernel support `noexec` parameter:

```
noexec          [X86]
            On X86-32 available only on PAE configured kernels.
            noexec=on: enable non-executable mappings (default)
            noexec=off: disable non-executable mappings
```

We can see it in the booting time:



After this we can see call of the:

```
    memblock_x86_reserve_range_setup_data();
```

function. This function is defined in the same arch/x86/kernel/setup.c source code file and remaps memory for the `setup_data` and reserved memory block for the `setup_data` (more about `setup_data` you can read in the previous part and about `ioremap` and `memblock` you can read in the Linux kernel memory management).

In the next step we can see following conditional statement:

```
    if (acpi_mps_check()) {
#ifdef CONFIG_X86_LOCAL_APIC
        disable_apic = 1;
#endif
        setup_clear_cpu_cap(X86_FEATURE_APIC);
    }
```

The first `acpi_mps_check` function from the arch/x86/kernel/acpi/boot.c depends on `CONFIG_X86_LOCAL_APIC` and `CONFIG_x86_MPPARSE` configuration options:

```
int __init acpi_mps_check(void)
{
#if defined(CONFIG_X86_LOCAL_APIC) && !defined(CONFIG_X86_MPPARSE)
        /* mptable code is not built-in*/
        if (acpi_disabled || acpi_noirq) {
                printk(KERN_WARNING "MPS support code is not built-in.\n"
                       "Using acpi=off or acpi=noirq or pci=noacpi "
                       "may have problem\n");
                 return 1;
        }
#endif
        return 0;
}
```

It checks the built-in `MPS` or MultiProcessor Specification table. If `CONFIG_X86_LOCAL_APIC` is set and `CONFIG_x86_MPPAARSE` is not set, `acpi_mps_check` prints warning message if the one of the command line options: `acpi=off` , `acpi=noirq` or `pci=noacpi` passed to the kernel. If `acpi_mps_check` returns `1` it means that we disable local APIC and clear `X86_FEATURE_APIC` bit in the of the current CPU with the `setup_clear_cpu_cap` macro. (more about CPU mask you can read in the CPU masks).

Early PCI dump

In the next step we make a dump of the PCI devices with the following code:

```
#ifdef CONFIG_PCI
    if (pci_early_dump_regs)
        early_dump_pci_devices();
#endif
```

`pci_early_dump_regs` variable defined in the arch/x86/pci/common.c and its value depends on the kernel command line parameter: `pci=earlydump` . We can find definition of this parameter in the drivers/pci/pci.c:

```
early_param("pci", pci_setup);
```

`pci_setup` function gets the string after the `pci=` and analyzes it. This function calls `pcibios_setup` which defined as `__weak` in the drivers/pci/pci.c and every architecture defines the same function which overrides `__weak` analog. For example `x86_64` architecture-dependent version is in the arch/x86/pci/common.c:

```
char *__init pcibios_setup(char *str) {
        ...
```

```
        ...
        ...
        } else if (!strcmp(str, "earlydump")) {
                pci_early_dump_regs = 1;
                return NULL;
        }
        ...
        ...
        ...
}
```

So, if `CONFIG_PCI` option is set and we passed `pci=earlydump` option to the kernel command line, next function which will be called - `early_dump_pci_devices` from the arch/x86/pci/early.c. This function checks `noearly` pci parameter with:

```
if (!early_pci_allowed())
        return;
```

and returns if it was passed. Each PCI domain can host up to `256` buses and each bus hosts up to 32 devices. So, we goes in a loop:

```
for (bus = 0; bus < 256; bus++) {
                for (slot = 0; slot < 32; slot++) {
                        for (func = 0; func < 8; func++) {
                        ...
                        ...
                        ...
                        }
                }
}
```

and read the `pci` config with the `read_pci_config` function.

That's all. We will not go deep in the `pci` details, but will see more details in the special `Drivers/PCI` part.

Finish with memory parsing

After the `early_dump_pci_devices`, there are a couple of function related with available memory and e820 which we collected in the First steps in the kernel setup part:

```
    /* update the e820_saved too */
    e820_reserve_setup_data();
    finish_e820_parsing();
    ...
    ...
    ...
    e820_add_kernel_range();
    trim_bios_range(void);
    max_pfn = e820_end_of_ram_pfn();
    early_reserve_e820_mpc_new();
```

Let's look on it. As you can see the first function is `e820_reserve_setup_data`. This function does almost the same as `memblock_x86_reserve_range_setup_data` which we saw above, but it also calls `e820_update_range` which adds new regions to the `e820map` with the given type which is `E820_RESERVED_KERN` in our case. The next function is `finish_e820_parsing` which sanitizes `e820map` with the `sanitize_e820_map` function. Besides this two functions we can see a couple of functions related to the e820. You can see it in the listing above. `e820_add_kernel_range` function takes the physical address of the kernel start and end:

```
u64 start = __pa_symbol(_text);
u64 size = __pa_symbol(_end) - start;
```

checks that `.text` `.data` and `.bss` marked as `E820RAM` in the `e820map` and prints the warning message if not. The next function `trm_bios_range` update first 4096 bytes in `e820Map` as `E820_RESERVED` and sanitizes it again with the call of the `sanitize_e820_map`. After this we get the last page frame number with the call of the `e820_end_of_ram_pfn` function. Every memory page has a unique number - `Page frame number` and `e820_end_of_ram_pfn` function returns the maximum with the call of the `e820_end_pfn`:

```
unsigned long __init e820_end_of_ram_pfn(void)
{
    return e820_end_pfn(MAX_ARCH_PFN);
}
```

where `e820_end_pfn` takes maximum page frame number on the certain architecture (`MAX_ARCH_PFN` is `0x400000000` for `x86_64`). In the `e820_end_pfn` we go through the all `e820` slots and check that `e820` entry has `E820_RAM` or `E820_PRAM` type because we calculate page frame numbers only for these types, gets the base address and end address of the page frame number for the current `e820` entry and makes some checks for these addresses:

```
for (i = 0; i < e820.nr_map; i++) {
        struct e820entry *ei = &e820.map[i];
        unsigned long start_pfn;
        unsigned long end_pfn;

        if (ei->type != E820_RAM && ei->type != E820_PRAM)
            continue;

        start_pfn = ei->addr >> PAGE_SHIFT;
        end_pfn = (ei->addr + ei->size) >> PAGE_SHIFT;

        if (start_pfn >= limit_pfn)
            continue;
        if (end_pfn > limit_pfn) {
            last_pfn = limit_pfn;
            break;
        }
        if (end_pfn > last_pfn)
            last_pfn = end_pfn;
}
```

```
    if (last_pfn > max_arch_pfn)
        last_pfn = max_arch_pfn;

    printk(KERN_INFO "e820: last_pfn = %#lx max_arch_pfn = %#lx\n",
             last_pfn, max_arch_pfn);
    return last_pfn;
```

After this we check that `last_pfn` which we got in the loop is not greater that maximum page frame number for the certain architecture (`x86_64` in our case), print information about last page frame number and return it. We can see the `last_pfn` in the `dmesg` output:

```
...
[    0.000000] e820: last_pfn = 0x41f000 max_arch_pfn = 0x400000000
...
```

After this, as we have calculated the biggest page frame number, we calculate `max_low_pfn` which is the biggest page frame number in the `low memory` or below first `4` gigabytes. If installed more than 4 gigabytes of RAM, `max_low_pfn` will be result of the `e820_end_of_low_ram_pfn` function which does the same `e820_end_of_ram_pfn` but with 4 gigabytes limit, in other way `max_low_pfn` will be the same as `max_pfn` :

```
if (max_pfn > (1UL<<(32 - PAGE_SHIFT)))
    max_low_pfn = e820_end_of_low_ram_pfn();
else
    max_low_pfn = max_pfn;

high_memory = (void *)__va(max_pfn * PAGE_SIZE - 1) + 1;
```

Next we calculate `high_memory` (defines the upper bound on direct map memory) with `__va` macro which returns a virtual address by the given physical memory.

DMI scanning

The next step after manipulations with different memory regions and `e820` slots is collecting information about computer. We will get all information with the Desktop Management Interface and following functions:

```
dmi_scan_machine();
dmi_memdev_walk();
```

First is `dmi_scan_machine` defined in the drivers/firmware/dmi_scan.c. This function goes through the System Management BIOS structures and extracts information. There are two ways specified to gain access to the `SMBIOS` table: get the pointer to the `SMBIOS` table from the EFI's configuration table and scanning the physical memory between `0xF0000` and `0x10000` addresses. Let's look on the second approach. `dmi_scan_machine` function remaps memory between `0xf0000` and `0x10000` with the `dmi_early_remap` which just expands to the `early_ioremap` :

```
void __init dmi_scan_machine(void)
{
    char __iomem *p, *q;
    char buf[32];
    ...
    ...
    ...
    p = dmi_early_remap(0xF0000, 0x10000);
    if (p == NULL)
            goto error;
```

and iterates over all `DMI` header address and find search `_SM_` string:

```
memset(buf, 0, 16);
for (q = p; q < p + 0x10000; q += 16) {
        memcpy_fromio(buf + 16, q, 16);
        if (!dmi_smbios3_present(buf) || !dmi_present(buf)) {
            dmi_available = 1;
            dmi_early_unmap(p, 0x10000);
            goto out;
        }
        memcpy(buf, buf + 16, 16);
}
```

`_SM_` string must be between `000F0000h` and `0x000FFFFF` . Here we copy 16 bytes to the `buf` with `memcpy_fromio` which is the same `memcpy` and execute `dmi_smbios3_present` and `dmi_present` on the buffer. These functions check that first 4 bytes is `_SM_` string, get `SMBIOS` version and gets `_DMI_` attributes as `DMI` structure table length, table address and etc... After one of these functions finish, you will see the result of it in the `dmesg` output:

```
[    0.000000] SMBIOS 2.7 present.
[    0.000000] DMI: Gigabyte Technology Co., Ltd. Z97X-UD5H-BK/Z97X-UD5H-BK, BIOS F6 06/17/2014
```

In the end of the `dmi_scan_machine`, we unmap the previously remapped memory:

```
dmi_early_unmap(p, 0x10000);
```

The second function is - `dmi_memdev_walk`. As you can understand it goes over memory devices. Let's look on it:

```
void __init dmi_memdev_walk(void)
{
    if (!dmi_available)
        return;

    if (dmi_walk_early(count_mem_devices) == 0 && dmi_memdev_nr) {
        dmi_memdev = dmi_alloc(sizeof(*dmi_memdev) * dmi_memdev_nr);
        if (dmi_memdev)
            dmi_walk_early(save_mem_devices);
    }
}
```

It checks that `DMI` available (we got it in the previous function - `dmi_scan_machine`) and collects information about memory devices with `dmi_walk_early` and `dmi_alloc` which defined as:

```
#ifdef CONFIG_DMI
RESERVE_BRK(dmi_alloc, 65536);
#endif
```

`RESERVE_BRK` defined in the arch/x86/include/asm/setup.h and reserves space with given size in the `brk` section.

```
init_hypervisor_platform();
x86_init.resources.probe_roms();
insert_resource(&iomem_resource, &code_resource);
insert_resource(&iomem_resource, &data_resource);
insert_resource(&iomem_resource, &bss_resource);
early_gart_iommu_check();
```

SMP config

The next step is parsing of the SMP configuration. We do it with the call of the `find_smp_config` function which just calls function:

```
static inline void find_smp_config(void)
{
        x86_init.mpparse.find_smp_config();
}
```

inside. `x86_init.mpparse.find_smp_config` is the `default_find_smp_config` function from the arch/x86/kernel/mpparse.c. In the `default_find_smp_config` function we are scanning a couple of memory regions for `SMP` config and return if they are found:

```
if (smp_scan_config(0x0, 0x400) ||
            smp_scan_config(639 * 0x400, 0x400) ||
            smp_scan_config(0xF0000, 0x10000))
            return;
```

First of all `smp_scan_config` function defines a couple of variables:

```
unsigned int *bp = phys_to_virt(base);
struct mpf_intel *mpf;
```

First is virtual address of the memory region where we will scan `SMP` config, second is the pointer to the `mpf_intel` structure. Let's try to understand what is it `mpf_intel` . All information stores in the multiprocessor configuration data structure. `mpf_intel` presents this structure and looks:

```
struct mpf_intel {
        char signature[4];
        unsigned int physptr;
        unsigned char length;
        unsigned char specification;
        unsigned char checksum;
        unsigned char feature1;
        unsigned char feature2;
        unsigned char feature3;
        unsigned char feature4;
        unsigned char feature5;
};
```

As we can read in the documentation - one of the main functions of the system BIOS is to construct the MP floating pointer structure and the MP configuration table. And operating system must have access to this information about the multiprocessor configuration and `mpf_intel` stores the physical address (look at second parameter) of the multiprocessor configuration table. So, `smp_scan_config` going in a loop through the given memory range and tries to find `MP floating pointer structure` there. It checks that current byte points to the `SMP` signature, checks checksum, checks if `mpf->specification` is 1 or 4(it must be `1` or `4` by specification) in the loop:

```
while (length > 0) {
if ((*bp == SMP_MAGIC_IDENT) &&
    (mpf->length == 1) &&
    !mpf_checksum((unsigned char *)bp, 16) &&
    ((mpf->specification == 1)
    || (mpf->specification == 4))) {

        mem = virt_to_phys(mpf);
        memblock_reserve(mem, sizeof(*mpf));
        if (mpf->physptr)
            smp_reserve_memory(mpf);
    }
}
```

reserves given memory block if search is successful with `memblock_reserve` and reserves physical address of the multiprocessor configuration table. You can find documentation about this in the - MultiProcessor Specification. You can read More details in the special part about `SMP` .

Additional early memory initialization routines

In the next step of the `setup_arch` we can see the call of the `early_alloc_pgt_buf` function which allocates the page table buffer for early stage. The page table buffer will be placed in the `brk` area. Let's look on its implementation:

```
void  __init early_alloc_pgt_buf(void)
{
        unsigned long tables = INIT_PGT_BUF_SIZE;
        phys_addr_t base;

        base = __pa(extend_brk(tables, PAGE_SIZE));

        pgt_buf_start = base >> PAGE_SHIFT;
        pgt_buf_end = pgt_buf_start;
        pgt_buf_top = pgt_buf_start + (tables >> PAGE_SHIFT);
}
```

First of all it get the size of the page table buffer, it will be `INIT_PGT_BUF_SIZE` which is `(6 * PAGE_SIZE)` in the current linux kernel 4.0. As we got the size of the page table buffer, we call `extend_brk` function with two parameters: size and align. As you can understand from its name, this function extends the `brk` area. As we can see in the linux kernel linker script `brk` is in memory right after the BSS:

```
    . = ALIGN(PAGE_SIZE);
    .brk : AT(ADDR(.brk) - LOAD_OFFSET) {
        __brk_base = .;
        . += 64 * 1024;        /* 64k alignment slop space */
        *(.brk_reservation)    /* areas brk users have reserved */
        __brk_limit = .;
    }
```

Or we can find it with `readelf` util:

```
  [25] .bss              NOBITS           ffffffff8199d000  00d9d000
       00000000000b4000  0000000000000000  WA       0     0     4096
  [26] .brk              NOBITS           ffffffff81a51000  00d9d000
       0000000000026000  0000000000000000  WA       0     0     1
```

After that we got physical address of the new `brk` with the `__pa` macro, we calculate the base address and the end of the page table buffer. In the next step as we got page table buffer, we reserve memory block for the brk area with the `reserve_brk` function:

```
static void __init reserve_brk(void)
{
    if (_brk_end > _brk_start)
        memblock_reserve(__pa_symbol(_brk_start),
                 _brk_end - _brk_start);

    _brk_start = 0;
}
```

Note that in the end of the `reserve_brk`, we set `brk_start` to zero, because after this we will not allocate it anymore. The next step after reserving memory block for the `brk`, we need to unmap out-of-range memory areas in the kernel mapping with the `cleanup_highmap` function. Remember that kernel mapping is `__START_KERNEL_map` and `_end - _text` or `level2_kernel_pgt` maps the kernel `_text`, `data` and `bss`. In the start of the `clean_high_map` we define these parameters:

```
unsigned long vaddr = __START_KERNEL_map;
unsigned long end = roundup((unsigned long)_end, PMD_SIZE) - 1;
pmd_t *pmd = level2_kernel_pgt;
pmd_t *last_pmd = pmd + PTRS_PER_PMD;
```

Now, as we defined start and end of the kernel mapping, we go in the loop through the all kernel page middle directory entries and clean entries which are not between `_text` and `end`:

```
for (; pmd < last_pmd; pmd++, vaddr += PMD_SIZE) {
        if (pmd_none(*pmd))
            continue;
        if (vaddr < (unsigned long) _text || vaddr > end)
            set_pmd(pmd, __pmd(0));
}
```

After this we set the limit for the `memblock` allocation with the `memblock_set_current_limit` function (read more about `memblock` you can in the Linux kernel memory management Part 2), it will be `ISA_END_ADDRESS` or `0x100000` and fill the `memblock` information according to `e820` with the call of the `memblock_x86_fill` function. You can see the result of this function in the kernel initialization time:

```
MEMBLOCK configuration:
 memory size = 0x1fff7ec00 reserved size = 0x1e30000
 memory.cnt  = 0x3
 memory[0x0]    [0x00000000001000-0x0000000009efff], 0x9e000 bytes flags: 0x0
 memory[0x1]    [0x00000000100000-0x000000bffdffff], 0xbfee0000 bytes flags: 0x0
 memory[0x2]    [0x00000100000000-0x0000023fffffff], 0x140000000 bytes flags: 0x0
 reserved.cnt  = 0x3
 reserved[0x0]    [0x0000000009f000-0x000000000fffff], 0x61000 bytes flags: 0x0
 reserved[0x1]    [0x00000001000000-0x00000001a57fff], 0xa58000 bytes flags: 0x0
 reserved[0x2]    [0x0000007ec89000-0x0000007fffffff], 0x1377000 bytes flags: 0x0
```

The rest functions after the `memblock_x86_fill` are: `early_reserve_e820_mpc_new` allocates additional slots in the `e820map` for MultiProcessor Specification table, `reserve_real_mode` - reserves low memory from `0x0` to 1 megabyte for the trampoline to the real mode (for rebooting, etc.), `trim_platform_memory_ranges` - trims certain memory regions started from `0x20050000`, `0x20110000`, etc. these regions must be excluded because Sandy Bridge has problems with these regions, `trim_low_memory_range` reserves the first 4 kilobyte page in `memblock`, `init_mem_mapping` function reconstructs direct memory mapping and setups the direct mapping of the physical memory at `PAGE_OFFSET`, `early_trap_pf_init` setups `#PF` handler (we will look on it in the chapter about interrupts) and `setup_real_mode` function setups trampoline to the real mode code.

That's all. You can note that this part will not cover all functions which are in the `setup_arch` (like `early_gart_iommu_check`, mtrr initialization, etc.). As I already wrote many times, `setup_arch` is big, and linux kernel is big. That's why I can't cover every line in the linux kernel. I don't think that we missed something important, but you can say something like: each line of code is important. Yes, it's true, but I missed them anyway, because I think that it is not realistic to cover full linux kernel. Anyway we will often return to the idea that we have already seen, and if something is unfamiliar, we will cover this theme.

Conclusion

It is the end of the sixth part about linux kernel initialization process. In this part we continued to dive in the `setup_arch` function again and it was long part, but we are not finished with it. Yes, `setup_arch` is big, hope that next part will be the last part about this function.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- MultiProcessor Specification
- NX bit
- Documentation/kernel-parameters.txt

- APIC
- CPU masks
- Linux kernel memory management
- PCI
- e820
- System Management BIOS
- System Management BIOS
- EFI
- SMP
- MultiProcessor Specification
- BSS
- SMBIOS specification
- Previous part

Kernel initialization. Part 7.

The End of the architecture-specific initialization, almost...

This is the seventh part of the Linux Kernel initialization process which covers insides of the `setup_arch` function from the arch/x86/kernel/setup.c. As you can know from the previous parts, the `setup_arch` function does some architecture-specific (in our case it is x86_64) initialization stuff like reserving memory for kernel code/data/bss, early scanning of the Desktop Management Interface, early dump of the PCI device and many many more. If you have read the previous part, you can remember that we've finished it at the `setup_real_mode` function. In the next step, as we set limit of the memblock to the all mapped pages, we can see the call of the `setup_log_buf` function from the kernel/printk/printk.c.

The `setup_log_buf` function setups kernel cyclic buffer and its length depends on the `CONFIG_LOG_BUF_SHIFT` configuration option. As we can read from the documentation of the `CONFIG_LOG_BUF_SHIFT` it can be between `12` and `21`. In the insides, buffer defined as array of chars:

```
#define __LOG_BUF_LEN (1 << CONFIG_LOG_BUF_SHIFT)
static char __log_buf[__LOG_BUF_LEN] __aligned(LOG_ALIGN);
static char *log_buf = __log_buf;
```

Now let's look on the implementation of the `setup_log_buf` function. It starts with check that current buffer is empty (It must be empty, because we just setup it) and another check that it is early setup. If setup of the kernel log buffer is not early, we call the `log_buf_add_cpu` function which increase size of the buffer for every CPU:

```
if (log_buf != __log_buf)
    return;

if (!early && !new_log_buf_len)
    log_buf_add_cpu();
```

We will not research `log_buf_add_cpu` function, because as you can see in the `setup_arch`, we call `setup_log_buf` as:

```
setup_log_buf(1);
```

where `1` means that it is early setup. In the next step we check `new_log_buf_len` variable which is updated length of the kernel log buffer and allocate new space for the buffer with the `memblock_virt_alloc` function for it, or just return.

As kernel log buffer is ready, the next function is `reserve_initrd`. You can remember that we already called the `early_reserve_initrd` function in the fourth part of the Kernel initialization. Now, as we reconstructed direct memory mapping in the `init_mem_mapping` function, we need to move initrd into directly mapped memory. The `reserve_initrd` function starts from the definition of the base address and end address of the `initrd` and check that `initrd` is provided by a bootloader. All the same as what we saw in the `early_reserve_initrd`. But instead of the reserving place in the `memblock` area with the call of the `memblock_reserve` function, we get the mapped size of the direct memory area and check that the size of the `initrd` is not greater than this area with:

```
mapped_size = memblock_mem_size(max_pfn_mapped);
if (ramdisk_size >= (mapped_size>>1))
    panic("initrd too large to handle, "
          "disabling initrd (%lld needed, %lld available)\n",
          ramdisk_size, mapped_size>>1);
```

You can see here that we call `memblock_mem_size` function and pass the `max_pfn_mapped` to it, where `max_pfn_mapped` contains the highest direct mapped page frame number. If you do not remember what is `page frame number`, explanation is simple: First `12` bits of the virtual address represent offset in the physical page or page frame. If we right-shift out `12` bits of the virtual address, we'll discard offset part and will get `Page Frame Number`. In the `memblock_mem_size` we go through the all memblock `mem` (not reserved) regions and calculates size of the mapped pages and return it to the `mapped_size` variable (see code above). As we got amount of the direct mapped memory, we check that size of the `initrd` is not greater than mapped pages. If it is greater we just call `panic` which halts the system and prints famous Kernel panic message. In the next step we print information about the `initrd` size. We can see the result of this in the `dmesg` output:

```
[0.000000] RAMDISK: [mem 0x36d20000-0x37687fff]
```

and relocate `initrd` to the direct mapping area with the `relocate_initrd` function. In the start of the `relocate_initrd` function we try to find a free area with the `memblock_find_in_range` function:

```
relocated_ramdisk = memblock_find_in_range(0, PFN_PHYS(max_pfn_mapped), area_size, PAGE_SIZE);

if (!relocated_ramdisk)
    panic("Cannot find place for new RAMDISK of size %lld\n",
           ramdisk_size);
```

The `memblock_find_in_range` function tries to find a free area in a given range, in our case from `0` to the maximum mapped physical address and size must equal to the aligned size of the `initrd`. If we didn't find a area with the given size, we call `panic` again. If all is good, we start to relocated RAM disk to the down of the directly mapped memory in the next step.

In the end of the `reserve_initrd` function, we free memblock memory which occupied by the ramdisk with the call of the:

```
memblock_free(ramdisk_image, ramdisk_end - ramdisk_image);
```

After we relocated `initrd` ramdisk image, the next function is `vsmp_init` from the arch/x86/kernel/vsmp_64.c. This function initializes support of the `ScaleMP vSMP`. As I already wrote in the previous parts, this chapter will not cover non-related `x86_64` initialization parts (for example as the current or `ACPI`, etc.). So we will skip implementation of this for now and will back to it in the part which cover techniques of parallel computing.

The next function is `io_delay_init` from the arch/x86/kernel/io_delay.c. This function allows to override default I/O delay `0x80` port. We already saw I/O delay in the Last preparation before transition into protected mode, now let's look on the `io_delay_init` implementation:

```
void __init io_delay_init(void)
{
    if (!io_delay_override)
        dmi_check_system(io_delay_0xed_port_dmi_table);
}
```

This function check `io_delay_override` variable and overrides I/O delay port if `io_delay_override` is set. We can set `io_delay_override` variably by passing `io_delay` option to the kernel command line. As we can read from the Documentation/kernel-parameters.txt, `io_delay` option is:

```
io_delay=    [X86] I/O delay method
    0x80
        Standard port 0x80 based delay
    0xed
        Alternate port 0xed based delay (needed on some systems)
    udelay
        Simple two microseconds delay
    none
        No delay
```

We can see `io_delay` command line parameter setup with the `early_param` macro in the arch/x86/kernel/io_delay.c

```
early_param("io_delay", io_delay_param);
```

More about `early_param` you can read in the previous part. So the `io_delay_param` function which setups `io_delay_override` variable will be called in the do_early_param function. `io_delay_param` function gets the argument of the `io_delay` kernel command line parameter and sets `io_delay_type` depends on it:

```
static int __init io_delay_param(char *s)
{
        if (!s)
                return -EINVAL;

        if (!strcmp(s, "0x80"))
                io_delay_type = CONFIG_IO_DELAY_TYPE_0X80;
        else if (!strcmp(s, "0xed"))
                io_delay_type = CONFIG_IO_DELAY_TYPE_0XED;
        else if (!strcmp(s, "udelay"))
                io_delay_type = CONFIG_IO_DELAY_TYPE_UDELAY;
        else if (!strcmp(s, "none"))
                io_delay_type = CONFIG_IO_DELAY_TYPE_NONE;
        else
                return -EINVAL;

        io_delay_override = 1;
        return 0;
}
```

The next functions are `acpi_boot_table_init`, `early_acpi_boot_init` and `initmem_init` after the `io_delay_init`, but as I wrote above we will not cover ACPI related stuff in this `Linux Kernel initialization process` chapter.

Allocate area for DMA

In the next step we need to allocate area for the Direct memory access with the `dma_contiguous_reserve` function which is defined in the drivers/base/dma-contiguous.c. `DMA` is a special mode when devices communicate with memory without CPU. Note that we pass one parameter - `max_pfn_mapped << PAGE_SHIFT`, to the `dma_contiguous_reserve` function and as you can understand from this expression, this is limit of the reserved memory. Let's look on the implementation of this function. It starts from the definition of the following variables:

```
phys_addr_t selected_size = 0;
phys_addr_t selected_base = 0;
phys_addr_t selected_limit = limit;
bool fixed = false;
```

where first represents size in bytes of the reserved area, second is base address of the reserved area, third is end address of the reserved area and the last `fixed` parameter shows where to place reserved area. If `fixed` is `1` we just reserve area with the `memblock_reserve`, if it is `0` we allocate space with the `kmemleak_alloc`. In the next step we check `size_cmdline` variable and if it is not equal to `-1` we fill all variables which you can see above with the values from the `cma` kernel command line parameter:

```
if (size_cmdline != -1) {
   ...
   ...
   ...
}
```

You can find in this source code file definition of the early parameter:

```
early_param("cma", early_cma);
```

where `cma` is:

```
cma=nn[MG]@[start[MG][-end[MG]]]
        [ARM,X86,KNL]
        Sets the size of kernel global memory area for
        contiguous memory allocations and optionally the
        placement constraint by the physical address range of
        memory allocations. A value of 0 disables CMA
        altogether. For more information, see
        include/linux/dma-contiguous.h
```

If we will not pass `cma` option to the kernel command line, `size_cmdline` will be equal to `-1`. In this way we need to calculate size of the reserved area which depends on the following kernel configuration options:

- `CONFIG_CMA_SIZE_SEL_MBYTES` - size in megabytes, default global `CMA` area, which is equal to `CMA_SIZE_MBYTES * SZ_1M` or `CONFIG_CMA_SIZE_MBYTES * 1M`;
- `CONFIG_CMA_SIZE_SEL_PERCENTAGE` - percentage of total memory;
- `CONFIG_CMA_SIZE_SEL_MIN` - use lower value;
- `CONFIG_CMA_SIZE_SEL_MAX` - use higher value.

As we calculated the size of the reserved area, we reserve area with the call of the `dma_contiguous_reserve_area` function which first of all calls:

```
ret = cma_declare_contiguous(base, size, limit, 0, 0, fixed, res_cma);
```

function. The `cma_declare_contiguous` reserves contiguous area from the given base address with given size. After we reserved area for the `DMA`, next function is the `memblock_find_dma_reserve`. As you can understand from its name, this function counts the reserved pages in the `DMA` area. This part will not cover all details of the `CMA` and `DMA`, because they are big. We will see much more details in the special part in the Linux Kernel Memory management which covers contiguous memory allocators and areas.

Initialization of the sparse memory

The next step is the call of the function - `x86_init.paging.pagetable_init`. If you try to find this function in the linux kernel source code, in the end of your search, you will see the following macro:

```
#define native_pagetable_init        paging_init
```

which expands as you can see to the call of the `paging_init` function from the arch/x86/mm/init_64.c. The `paging_init` function initializes sparse memory and zone sizes. First of all what's zones and what is it `Sparsemem`. The `Sparsemem` is a special foundation in the linux kernel memory manager which used to split memory area into different memory banks in the NUMA systems. Let's look on the implementation of the `paginig_init` function:

```
void __init paging_init(void)
{
        sparse_memory_present_with_active_regions(MAX_NUMNODES);
        sparse_init();

        node_clear_state(0, N_MEMORY);
        if (N_MEMORY != N_NORMAL_MEMORY)
                node_clear_state(0, N_NORMAL_MEMORY);
```

```
        zone_sizes_init();
}
```

As you can see there is call of the `sparse_memory_present_with_active_regions` function which records a memory area for every `NUMA` node to the array of the `mem_section` structure which contains a pointer to the structure of the array of `struct page`. The next `sparse_init` function allocates non-linear `mem_section` and `mem_map`. In the next step we clear state of the movable memory nodes and initialize sizes of zones. Every `NUMA` node is divided into a number of pieces which are called - `zones`. So, `zone_sizes_init` function from the arch/x86/mm/init.c initializes size of zones.

Again, this part and next parts do not cover this theme in full details. There will be special part about `NUMA`.

vsyscall mapping

The next step after `SparseMem` initialization is setting of the `trampoline_cr4_features` which must contain content of the `cr4` Control register. First of all we need to check that current CPU has support of the `cr4` register and if it has, we save its content to the `trampoline_cr4_features` which is storage for `cr4` in the real mode:

```
if (boot_cpu_data.cpuid_level >= 0) {
    mmu_cr4_features = __read_cr4();
    if (trampoline_cr4_features)
        *trampoline_cr4_features = mmu_cr4_features;
}
```

The next function which you can see is `map_vsyscal` from the arch/x86/kernel/vsyscall_64.c. This function maps memory space for vsyscalls and depends on `CONFIG_X86_VSYSCALL_EMULATION` kernel configuration option. Actually `vsyscall` is a special segment which provides fast access to the certain system calls like `getcpu`, etc. Let's look on implementation of this function:

```
void __init map_vsyscall(void)
{
        extern char __vsyscall_page;
        unsigned long physaddr_vsyscall = __pa_symbol(&__vsyscall_page);

        if (vsyscall_mode != NONE)
                __set_fixmap(VSYSCALL_PAGE, physaddr_vsyscall,
                             vsyscall_mode == NATIVE
                             ? PAGE_KERNEL_VSYSCALL
                             : PAGE_KERNEL_VVAR);

        BUILD_BUG_ON((unsigned long)__fix_to_virt(VSYSCALL_PAGE) !=
                     (unsigned long)VSYSCALL_ADDR);
}
```

In the beginning of the `map_vsyscall` we can see definition of two variables. The first is extern variable `__vsyscall_page`. As a extern variable, it defined somewhere in other source code file. Actually we can see definition of the `__vsyscall_page` in the arch/x86/kernel/vsyscall_emu_64.S. The `__vsyscall_page` symbol points to the aligned calls of the `vsyscalls` as `gettimeofday`, etc.:

```
    .globl __vsyscall_page
    .balign PAGE_SIZE, 0xcc
    .type __vsyscall_page, @object
__vsyscall_page:

    mov $__NR_gettimeofday, %rax
    syscall
    ret

    .balign 1024, 0xcc
```

```
    mov $__NR_time, %rax
    syscall
    ret
    ...
    ...
    ...
```

The second variable is `physaddr_vsyscall` which just stores physical address of the `__vsyscall_page` symbol. In the next step we check the `vsyscall_mode` variable, and if it is not equal to `NONE` , it is `EMULATE` by default:

```
static enum { EMULATE, NATIVE, NONE } vsyscall_mode = EMULATE;
```

And after this check we can see the call of the `__set_fixmap` function which calls `native_set_fixmap` with the same parameters:

```
void native_set_fixmap(enum fixed_addresses idx, unsigned long phys, pgprot_t flags)
{
        __native_set_fixmap(idx, pfn_pte(phys >> PAGE_SHIFT, flags));
}

void __native_set_fixmap(enum fixed_addresses idx, pte_t pte)
{
        unsigned long address = __fix_to_virt(idx);

        if (idx >= __end_of_fixed_addresses) {
                BUG();
                return;
        }
        set_pte_vaddr(address, pte);
        fixmaps_set++;
}
```

Here we can see that `native_set_fixmap` makes value of `Page Table Entry` from the given physical address (physical address of the `__vsyscall_page` symbol in our case) and calls internal function - `__native_set_fixmap` . Internal function gets the virtual address of the given `fixed_addresses` index (`VSYSCALL_PAGE` in our case) and checks that given index is not greater than end of the fix-mapped addresses. After this we set page table entry with the call of the `set_pte_vaddr` function and increase count of the fix-mapped addresses. And in the end of the `map_vsyscall` we check that virtual address of the `VSYSCALL_PAGE` (which is first index in the `fixed_addresses`) is not greater than `VSYSCALL_ADDR` which is `-10UL << 20` or `ffffffffff600000` with the `BUILD_BUG_ON` macro:

```
BUILD_BUG_ON((unsigned long)__fix_to_virt(VSYSCALL_PAGE) !=
                     (unsigned long)VSYSCALL_ADDR);
```

Now `vsyscall` area is in the `fix-mapped` area. That's all about `map_vsyscall` , if you do not know anything about fix-mapped addresses, you can read Fix-Mapped Addresses and ioremap. We will see more about `vsyscalls` in the `vsyscalls and vdso` part.

Getting the SMP configuration

You may remember how we made a search of the SMP configuration in the previous part. Now we need to get the `SMP` configuration if we found it. For this we check `smp_found_config` variable which we set in the `smp_scan_config` function (read about it the previous part) and call the `get_smp_config` function:

```
if (smp_found_config)
    get_smp_config();
```

The `get_smp_config` expands to the `x86_init.mpparse.default_get_smp_config` function which is defined in the arch/x86/kernel/mpparse.c. This function defines a pointer to the multiprocessor floating pointer structure - `mpf_intel` (you can read about it in the previous part) and does some checks:

```
struct mpf_intel *mpf = mpf_found;

if (!mpf)
    return;

if (acpi_lapic && early)
   return;
```

Here we can see that multiprocessor configuration was found in the `smp_scan_config` function or just return from the function if not. The next check is `acpi_lapic` and `early`. And as we did this checks, we start to read the `SMP` configuration. As we finished reading it, the next step is - `prefill_possible_map` function which makes preliminary filling of the possible CPU's `cpumask` (more about it you can read in the Introduction to the cpumasks).

The rest of the setup_arch

Here we are getting to the end of the `setup_arch` function. The rest of function of course is important, but details about these stuff will not will not be included in this part. We will just take a short look on these functions, because although they are important as I wrote above, but they cover non-generic kernel features related with the `NUMA`, `SMP`, `ACPI` and `APICs`, etc. First of all, the next call of the `init_apic_mappings` function. As we can understand this function sets the address of the local APIC. The next is `x86_io_apic_ops.init` and this function initializes I/O APIC. Please note that we will see all details related with `APIC` in the chapter about interrupts and exceptions handling. In the next step we reserve standard I/O resources like `DMA`, `TIMER`, `FPU`, etc., with the call of the `x86_init.resources.reserve_resources` function. Following is `mcheck_init` function initializes `Machine check Exception` and the last is `register_refined_jiffies` which registers jiffy (There will be separate chapter about timers in the kernel).

So that's all. Finally we have finished with the big `setup_arch` function in this part. Of course as I already wrote many times, we did not see full details about this function, but do not worry about it. We will be back more than once to this function from different chapters for understanding how different platform-dependent parts are initialized.

That's all, and now we can back to the `start_kernel` from the `setup_arch`.

Back to the main.c

As I wrote above, we have finished with the `setup_arch` function and now we can back to the `start_kernel` function from the init/main.c. As you may remember or saw yourself, `start_kernel` function as big as the `setup_arch`. So the couple of the next part will be dedicated to learning of this function. So, let's continue with it. After the `setup_arch` we can see the call of the `mm_init_cpumask` function. This function sets the cpumask pointer to the memory descriptor `cpumask`. We can look on its implementation:

```
static inline void mm_init_cpumask(struct mm_struct *mm)
{
#ifdef CONFIG_CPUMASK_OFFSTACK
        mm->cpu_vm_mask_var = &mm->cpumask_allocation;
#endif
        cpumask_clear(mm->cpu_vm_mask_var);
}
```

As you can see in the init/main.c, we pass memory descriptor of the init process to the `mm_init_cpumask` and depends on `CONFIG_CPUMASK_OFFSTACK` configuration option we clear TLB switch `cpumask`.

In the next step we can see the call of the following function:

```
setup_command_line(command_line);
```

This function takes pointer to the kernel command line allocates a couple of buffers to store command line. We need a couple of buffers, because one buffer used for future reference and accessing to command line and one for parameter parsing. We will allocate space for the following buffers:

* `saved_command_line` - will contain boot command line;
* `initcall_command_line` - will contain boot command line. will be used in the `do_initcall_level` ;
* `static_command_line` - will contain command line for parameters parsing.

We will allocate space with the `memblock_virt_alloc` function. This function calls `memblock_virt_alloc_try_nid` which allocates boot memory block with `memblock_reserve` if slab is not available or uses `kzalloc_node` (more about it will be in the linux memory management chapter). The `memblock_virt_alloc` uses `BOOTMEM_LOW_LIMIT` (physical address of the `(PAGE_OFFSET + 0x1000000)` value) and `BOOTMEM_ALLOC_ACCESSIBLE` (equal to the current value of the `memblock.current_limit`) as minimum address of the memory region and maximum address of the memory region.

Let's look on the implementation of the `setup_command_line` :

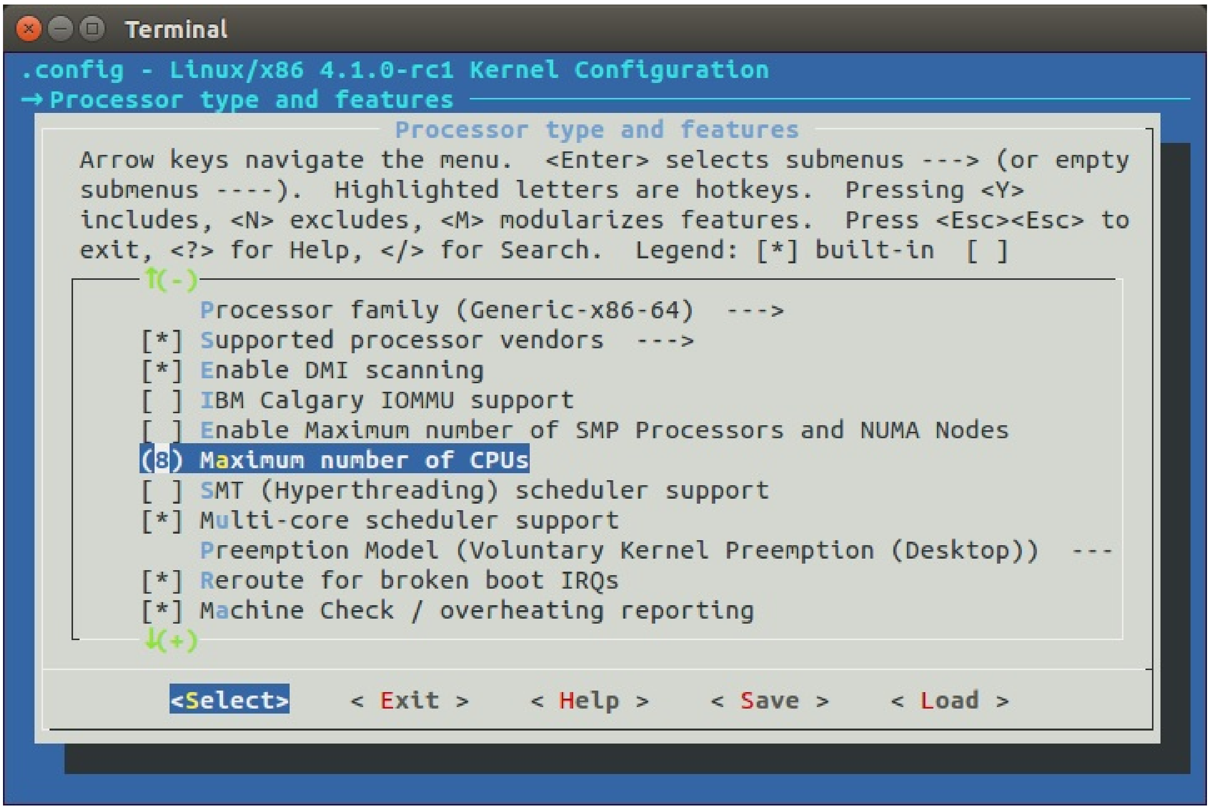
```
static void __init setup_command_line(char *command_line)
{
        saved_command_line =
                memblock_virt_alloc(strlen(boot_command_line) + 1, 0);
        initcall_command_line =
                memblock_virt_alloc(strlen(boot_command_line) + 1, 0);
        static_command_line = memblock_virt_alloc(strlen(command_line) + 1, 0);
        strcpy(saved_command_line, boot_command_line);
        strcpy(static_command_line, command_line);
 }
```

Here we can see that we allocate space for the three buffers which will contain kernel command line for the different purposes (read above). And as we allocated space, we store `boot_command_line` in the `saved_command_line` and `command_line` (kernel command line from the `setup_arch`) to the `static_command_line` .

The next function after the `setup_command_line` is the `setup_nr_cpu_ids` . This function setting `nr_cpu_ids` (number of CPUs) according to the last bit in the `cpu_possible_mask` (more about it you can read in the chapter describes cpumasks concept). Let's look on its implementation:

```
void __init setup_nr_cpu_ids(void)
{
        nr_cpu_ids = find_last_bit(cpumask_bits(cpu_possible_mask),NR_CPUS) + 1;
}
```

Here `nr_cpu_ids` represents number of CPUs, `NR_CPUS` represents the maximum number of CPUs which we can set in configuration time:



Actually we need to call this function, because `NR_CPUS` can be greater than actual amount of the CPUs in the your computer. Here we can see that we call `find_last_bit` function and pass two parameters to it:

- `cpu_possible_mask` bits;
- maximum number of CPUS.

In the `setup_arch` we can find the call of the `prefill_possible_map` function which calculates and writes to the `cpu_possible_mask` actual number of the CPUs. We call the `find_last_bit` function which takes the address and maximum size to search and returns bit number of the first set bit. We passed `cpu_possible_mask` bits and maximum number of the CPUs. First of all the `find_last_bit` function splits given `unsigned long` address to the words:

```
words = size / BITS_PER_LONG;
```

where `BITS_PER_LONG` is `64` on the `x86_64` . As we got amount of words in the given size of the search data, we need to check is given size does not contain partial words with the following check:

```
if (size & (BITS_PER_LONG-1)) {
         tmp = (addr[words] & (~0UL >> (BITS_PER_LONG
                                  - (size & (BITS_PER_LONG-1)))));
         if (tmp)
                 goto found;
}
```

if it contains partial word, we mask the last word and check it. If the last word is not zero, it means that current word contains at least one set bit. We go to the `found` label:

```
found:
    return words * BITS_PER_LONG + __fls(tmp);
```

Here you can see `__fls` function which returns last set bit in a given word with help of the `bsr` instruction:

```
static inline unsigned long __fls(unsigned long word)
{
        asm("bsr %1,%0"
            : "=r" (word)
            : "rm" (word));
        return word;
}
```

The `bsr` instruction which scans the given operand for first bit set. If the last word is not partial we going through the all words in the given address and trying to find first set bit:

```
while (words) {
    tmp = addr[--words];
    if (tmp) {
found:
        return words * BITS_PER_LONG + __fls(tmp);
    }
}
```

Here we put the last word to the `tmp` variable and check that `tmp` contains at least one set bit. If a set bit found, we return the number of this bit. If no one words do not contains set bit we just return given size:

```
return size;
```

After this `nr_cpu_ids` will contain the correct amount of the available CPUs.

That's all.

Conclusion

It is the end of the seventh part about the linux kernel initialization process. In this part, finally we have finished with the `setup_arch` function and returned to the `start_kernel` function. In the next part we will continue to learn generic kernel code from the `start_kernel` and will continue our way to the first `init` process.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- Desktop Management Interface
- x86_64
- initrd
- Kernel panic
- Documentation/kernel-parameters.txt
- ACPI
- Direct memory access
- NUMA
- Control register
- vsyscalls
- SMP
- jiffy

- Previous part

Kernel initialization. Part 8.

Scheduler initialization

This is the eighth part of the Linux kernel initialization process chapter and we stopped on the `setup_nr_cpu_ids` function in the previous part.

The main point of this part is scheduler initialization. But before we will start to learn initialization process of the scheduler, we need to do some stuff. The next step in the init/main.c is the `setup_per_cpu_areas` function. This function setups memory areas for the `percpu` variables, more about it you can read in the special part about the Per-CPU variables. After `percpu` areas is up and running, the next step is the `smp_prepare_boot_cpu` function.

This function does some preparations for symmetric multiprocessing. Since this function is architecture specific, it is located in the arch/x86/include/asm/smp.h Linux kernel header file. Let's look at the definition of this function:

```
static inline void smp_prepare_boot_cpu(void)
{
         smp_ops.smp_prepare_boot_cpu();
}
```

We may see here that it just calls the `smp_prepare_boot_cpu` callback of the `smp_ops` structure. If we look at the definition of instance of this structure from the arch/x86/kernel/smp.c source code file, we will see that the `smp_prepare_boot_cpu` expands to the call of the `native_smp_prepare_boot_cpu` function:

```
struct smp_ops smp_ops = {
    ...
    ...
    ...
    smp_prepare_boot_cpu = native_smp_prepare_boot_cpu,
    ...
    ...
    ...
}
EXPORT_SYMBOL_GPL(smp_ops);
```

The `native_smp_prepare_boot_cpu` function looks:

```
void __init native_smp_prepare_boot_cpu(void)
{
        int me = smp_processor_id();
        switch_to_new_gdt(me);
        cpumask_set_cpu(me, cpu_callout_mask);
        per_cpu(cpu_state, me) = CPU_ONLINE;
}
```

and executes following things: first of all it gets the `id` of the current CPU (which is Bootstrap processor and its `id` is zero for this moment) with the `smp_processor_id` function. I will not explain how the `smp_processor_id` works, because we already saw it in the Kernel entry point part. After we've got processor `id` number we reload Global Descriptor Table for the given CPU with the `switch_to_new_gdt` function:

```
void switch_to_new_gdt(int cpu)
{
        struct desc_ptr gdt_descr;
```

```
        gdt_descr.address = (long)get_cpu_gdt_table(cpu);
        gdt_descr.size = GDT_SIZE - 1;
        load_gdt(&gdt_descr);
        load_percpu_segment(cpu);
}
```

The `gdt_descr` variable represents pointer to the `GDT` descriptor here (we already saw definition of a `desc_ptr` structure in the Early interrupt and exception handling part). We get the address and the size of the `GDT` descriptor for the `CPU` with the given `id`. The `GDT_SIZE` is `256` or:

```
#define GDT_SIZE (GDT_ENTRIES * 8)
```

and the address of the descriptor we will get with the `get_cpu_gdt_table`:

```
static inline struct desc_struct *get_cpu_gdt_table(unsigned int cpu)
{
        return per_cpu(gdt_page, cpu).gdt;
}
```

The `get_cpu_gdt_table` uses `per_cpu` macro for getting value of a `gdt_page` percpu variable for the given CPU number (bootstrap processor with `id` - 0 in our case).

You may ask the following question: so, if we can access `gdt_page` percpu variable, where it was defined? Actually we already saw it in this book. If you have read the first part of this chapter, you can remember that we saw definition of the `gdt_page` in the arch/x86/kernel/head_64.S:

```
early_gdt_descr:
    .word    GDT_ENTRIES*8-1
early_gdt_descr_base:
    .quad    INIT_PER_CPU_VAR(gdt_page)
```

and if we will look on the linker file we can see that it locates after the `__per_cpu_load` symbol:

```
#define INIT_PER_CPU(x) init_per_cpu__##x = x + __per_cpu_load
INIT_PER_CPU(gdt_page);
```

and filled `gdt_page` in the arch/x86/kernel/cpu/common.c:

```
DEFINE_PER_CPU_PAGE_ALIGNED(struct gdt_page, gdt_page) = { .gdt = {
#ifdef CONFIG_X86_64
    [GDT_ENTRY_KERNEL32_CS]        = GDT_ENTRY_INIT(0xc09b, 0, 0xfffff),
    [GDT_ENTRY_KERNEL_CS]        = GDT_ENTRY_INIT(0xa09b, 0, 0xfffff),
    [GDT_ENTRY_KERNEL_DS]        = GDT_ENTRY_INIT(0xc093, 0, 0xfffff),
    [GDT_ENTRY_DEFAULT_USER32_CS]    = GDT_ENTRY_INIT(0xc0fb, 0, 0xfffff),
    [GDT_ENTRY_DEFAULT_USER_DS]    = GDT_ENTRY_INIT(0xc0f3, 0, 0xfffff),
    [GDT_ENTRY_DEFAULT_USER_CS]    = GDT_ENTRY_INIT(0xa0fb, 0, 0xfffff),
    ...
    ...
    ...
```

more about `percpu` variables you can read in the Per-CPU variables part. As we got address and size of the `GDT` descriptor we reload `GDT` with the `load_gdt` which just execute `lgdt` instruct and load `percpu_segment` with the following function:

```
void load_percpu_segment(int cpu) {
    loadsegment(gs, 0);
    wrmsrl(MSR_GS_BASE, (unsigned long)per_cpu(irq_stack_union.gs_base, cpu));
    load_stack_canary_segment();
}
```

The base address of the `percpu` area must contain `gs` register (or `fs` register for `x86`), so we are using `loadsegment` macro and pass `gs` . In the next step we writes the base address if the IRQ stack and setup stack canary (this is only for `x86_32`). After we load new `GDT` , we fill `cpu_callout_mask` bitmap with the current cpu and set cpu state as online with the setting `cpu_state` percpu variable for the current processor - `CPU_ONLINE` :

```
cpumask_set_cpu(me, cpu_callout_mask);
per_cpu(cpu_state, me) = CPU_ONLINE;
```

So, what is `cpu_callout_mask` bitmap... As we initialized bootstrap processor (processor which is booted the first on `x86`) the other processors in a multiprocessor system are known as `secondary processors` . Linux kernel uses following two bitmasks:

- `cpu_callout_mask`
- `cpu_callin_mask`

After bootstrap processor initialized, it updates the `cpu_callout_mask` to indicate which secondary processor can be initialized next. All other or secondary processors can do some initialization stuff before and check the `cpu_callout_mask` on the boostrap processor bit. Only after the bootstrap processor filled the `cpu_callout_mask` with this secondary processor, it will continue the rest of its initialization. After that the certain processor finish its initialization process, the processor sets bit in the `cpu_callin_mask` . Once the bootstrap processor finds the bit in the `cpu_callin_mask` for the current secondary processor, this processor repeats the same procedure for initialization of one of the remaining secondary processors. In a short words it works as i described, but we will see more details in the chapter about `SMP` .

That's all. We did all `SMP` boot preparation.

Build zonelists

In the next step we can see the call of the `build_all_zonelists` function. This function sets up the order of zones that allocations are preferred from. What are zones and what's order we will understand soon. For the start let's see how linux kernel considers physical memory. Physical memory is split into banks which are called - `nodes` . If you has no hardware support for `NUMA` , you will see only one node:

```
$ cat /sys/devices/system/node/node0/numastat
numa_hit 72452442
numa_miss 0
numa_foreign 0
interleave_hit 12925
local_node 72452442
other_node 0
```

Every `node` is presented by the `struct pglist_data` in the linux kernel. Each node is divided into a number of special blocks which are called - `zones` . Every zone is presented by the `zone struct` in the linux kernel and has one of the type:

- `ZONE_DMA` - 0-16M;
- `ZONE_DMA32` - used for 32 bit devices that can only do DMA areas below 4G;
- `ZONE_NORMAL` - all RAM from the 4GB on the `x86_64` ;
- `ZONE_HIGHMEM` - absent on the `x86_64` ;
- `ZONE_MOVABLE` - zone which contains movable pages.

which are presented by the `zone_type` enum. We can get information about zones with the:

```
$ cat /proc/zoneinfo
Node 0, zone      DMA
  pages free     3975
        min      3
```

```
         low      3
         ...
         ...
Node 0, zone    DMA32
  pages free     694163
         min      875
         low      1093
         ...
         ...
Node 0, zone   Normal
  pages free     2529995
         min      3146
         low      3932
         ...
         ...
```

As I wrote above all nodes are described with the `pglist_data` or `pg_data_t` structure in memory. This structure is defined in the include/linux/mmzone.h. The `build_all_zonelists` function from the mm/page_alloc.c constructs an ordered `zonelist` (of different zones `DMA`, `DMA32`, `NORMAL`, `HIGH_MEMORY`, `MOVABLE`) which specifies the zones/nodes to visit when a selected `zone` or `node` cannot satisfy the allocation request. That's all. More about `NUMA` and multiprocessor systems will be in the special part.

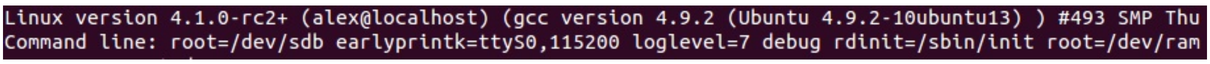
The rest of the stuff before scheduler initialization

Before we will start to dive into linux kernel scheduler initialization process we must do a couple of things. The first thing is the `page_alloc_init` function from the mm/page_alloc.c. This function looks pretty easy:

```
void __init page_alloc_init(void)
{
    int ret;

    ret = cpuhp_setup_state_nocalls(CPUHP_PAGE_ALLOC_DEAD,
                                    "mm/page_alloc:dead", NULL,
                                    page_alloc_cpu_dead);
    WARN_ON(ret < 0);
}
```

It setups setup the `startup` and `teardown` callbacks (second and third parameters) for the `CPUHP_PAGE_ALLOC_DEAD` cpu hotplug state. Of course the implementation of this function depends on the `CONFIG_HOTPLUG_CPU` kernel configuration option and if this option is set, such callbacks will be set for all cpu(s) in the system depends on their `hotplug` states. hotplug mechanism is a big theme and it will not be described in this book.

After this function we can see the kernel command line in the initialization output:



And a couple of functions such as `parse_early_param` and `parse_args` which handles linux kernel command line. You may remember that we already saw the call of the `parse_early_param` function in the sixth part of the kernel initialization chapter, so why we call it again? Answer is simple: we call this function in the architecture-specific code (`x86_64` in our case), but not all architecture calls this function. And we need to call the second function `parse_args` to parse and handle non-early command line arguments.

In the next step we can see the call of the `jump_label_init` from the kernel/jump_label.c. and initializes jump label.

After this we can see the call of the `setup_log_buf` function which setups the printk log buffer. We already saw this function in the seventh part of the linux kernel initialization process chapter.

PID hash initialization

The next is `pidhash_init` function. As you know each process has assigned a unique number which called - `process identification number` or `PID`. Each process generated with fork or clone is automatically assigned a new unique `PID` value by the kernel. The management of `PIDs` centered around the two special data structures: `struct pid` and `struct upid`. First structure represents information about a `PID` in the kernel. The second structure represents the information that is visible in a specific namespace. All `PID` instances stored in the special hash table:

```
static struct hlist_head *pid_hash;
```

This hash table is used to find the pid instance that belongs to a numeric `PID` value. So, `pidhash_init` initializes this hash table. In the start of the `pidhash_init` function we can see the call of the `alloc_large_system_hash`:

```
pid_hash = alloc_large_system_hash("PID", sizeof(*pid_hash), 0, 18,
                                   HASH_EARLY | HASH_SMALL,
                                   &pidhash_shift, NULL,
                                   0, 4096);
```

The number of elements of the `pid_hash` depends on the `RAM` configuration, but it can be between `2^4` and `2^12`. The `pidhash_init` computes the size and allocates the required storage (which is `hlist` in our case - the same as doubly linked list, but contains one pointer instead on the struct hlist_head]. The `alloc_large_system_hash` function allocates a large system hash table with `memblock_virt_alloc_nopanic` if we pass `HASH_EARLY` flag (as it in our case) or with `__vmalloc` if we did no pass this flag.

The result we can see in the `dmesg` output:

```
$ dmesg | grep hash
[    0.000000] PID hash table entries: 4096 (order: 3, 32768 bytes)
...
...
...
```

That's all. The rest of the stuff before scheduler initialization is the following functions: `vfs_caches_init_early` does early initialization of the virtual file system (more about it will be in the chapter which will describe virtual file system), `sort_main_extable` sorts the kernel's built-in exception table entries which are between `__start___ex_table` and `__stop___ex_table`, and `trap_init` initializes trap handlers (more about last two function we will know in the separate chapter about interrupts).

The last step before the scheduler initialization is initialization of the memory manager with the `mm_init` function from the init/main.c. As we can see, the `mm_init` function initializes different parts of the linux kernel memory manager:

```
page_ext_init_flatmem();
mem_init();
kmem_cache_init();
percpu_init_late();
pgtable_init();
vmalloc_init();
```

The first is `page_ext_init_flatmem` which depends on the `CONFIG_SPARSEMEM` kernel configuration option and initializes extended data per page handling. The `mem_init` releases all `bootmem`, the `kmem_cache_init` initializes kernel cache, the `percpu_init_late` - replaces `percpu` chunks with those allocated by slub, the `pgtable_init` - initializes the `page->ptl` kernel cache, the `vmalloc_init` - initializes `vmalloc`. Please, **NOTE** that we will not dive into details about all of these functions and concepts, but we will see all of they it in the Linux kernel memory manager chapter.

That's all. Now we can look on the `scheduler`.

Scheduler initialization

And now we come to the main purpose of this part - initialization of the task scheduler. I want to say again as I already did it many times, you will not see the full explanation of the scheduler here, there will be special separate chapter about this. Here will be described first initial scheduler mechanisms which are initialized first of all. So let's start.

Our current point is the `sched_init` function from the kernel/sched/core.c kernel source code file and as we can understand from the function's name, it initializes scheduler. Let's start to dive into this function and try to understand how the scheduler is initialized. At the start of the `sched_init` function we can see the following call:

```
sched_clock_init();
```

The `sched_clock_init` is pretty easy function and as we may see it just sets `sched_clock_init` variable:

```
void sched_clock_init(void)
{
	sched_clock_running = 1;
}
```

that will be used later. At the next step is initialization of the array of `waitqueues` :

```
for (i = 0; i < WAIT_TABLE_SIZE; i++)
	init_waitqueue_head(bit_wait_table + i);
```

where `bit_wait_table` is defined as:

```
#define WAIT_TABLE_BITS 8
#define WAIT_TABLE_SIZE (1 << WAIT_TABLE_BITS)
static wait_queue_head_t bit_wait_table[WAIT_TABLE_SIZE] __cacheline_aligned;
```

The `bit_wait_table` is array of wait queues that will be used for wait/wake up of processes depends on the value of a designated bit. The next step after initialization of `waitqueues` array is calculating size of memory to allocate for the `root_task_group` . As we may see this size depends on two following kernel configuration options:

```
#ifdef CONFIG_FAIR_GROUP_SCHED
         alloc_size += 2 * nr_cpu_ids * sizeof(void **);
#endif
#ifdef CONFIG_RT_GROUP_SCHED
         alloc_size += 2 * nr_cpu_ids * sizeof(void **);
#endif
```

- `CONFIG_FAIR_GROUP_SCHED` ;
- `CONFIG_RT_GROUP_SCHED` .

Both of these options provide two different planning models. As we can read from the documentation, the current scheduler - `CFS` or `Completely Fair Scheduler` use a simple concept. It models process scheduling as if the system has an ideal multitasking processor where each process would receive `1/n` processor time, where `n` is the number of the runnable processes. The scheduler uses the special set of rules. These rules determine when and how to select a new process to run and they are called `scheduling policy` .

The `Completely Fair Scheduler` supports following `normal` or in other words `non-real-time` scheduling policies:

- `SCHED_NORMAL` ;
- `SCHED_BATCH` ;
- `SCHED_IDLE` .

The `SCHED_NORMAL` is used for the most normal applications, the amount of cpu each process consumes is mostly determined by the nice value, the `SCHED_BATCH` used for the 100% non-interactive tasks and the `SCHED_IDLE` runs tasks only when the processor has no task to run besides this task.

The `real-time` policies are also supported for the time-critical applications: `SCHED_FIFO` and `SCHED_RR`. If you've read something about the Linux kernel scheduler, you can know that it is modular. It means that it supports different algorithms to schedule different types of processes. Usually this modularity is called `scheduler classes`. These modules encapsulate scheduling policy details and are handled by the scheduler core without knowing too much about them.

Now let's get back to the our code and look on the two configuration options: `CONFIG_FAIR_GROUP_SCHED` and `CONFIG_RT_GROUP_SCHED`. The least unit which scheduler operates is an individual task or thread. But a process is not only one type of entities of which the scheduler may operate. Both of these options provides support for group scheduling. The first one option provides support for group scheduling with `completely fair scheduler` policies and the second with `real-time` policies respectively.

In simple words, group scheduling is a feature that allows us to schedule a set of tasks as if a single task. For example, if you create a group with two tasks on the group, then this group is just like one normal task, from the kernel perspective. After a group is scheduled, the scheduler will pick a task from this group and it will be scheduled inside the group. So, such mechanism allows us to build hierarchies and manage their resources. Although a minimal unit of scheduling is a process, the Linux kernel scheduler does not use `task_struct` structure under the hood. There is special `sched_entity` structure that is used by the Linux kernel scheduler as scheduling unit.

So, the current goal is to calculate a space to allocate for a `sched_entity(ies)` of the root task group and we do it two times with:

```
#ifdef CONFIG_FAIR_GROUP_SCHED
         alloc_size += 2 * nr_cpu_ids * sizeof(void **);
#endif
#ifdef CONFIG_RT_GROUP_SCHED
         alloc_size += 2 * nr_cpu_ids * sizeof(void **);
#endif
```

The first is for case when scheduling of task groups is enabled with `completely fair` scheduler and the second is for the same purpose by in a case of `real-time` scheduler. So here we calculate size which is equal to size of a pointer multiplied on amount of CPUs in the system and multiplied to `2`. We need to multiply this on `2` as we will need to allocate a space for two things:

- scheduler entity structure;
- `runqueue`.

After we have calculated size, we allocate a space with the `kzalloc` function and set pointers of `sched_entity` and `runquques` there:

```
ptr = (unsigned long)kzalloc(alloc_size, GFP_NOWAIT);

#ifdef CONFIG_FAIR_GROUP_SCHED
        root_task_group.se = (struct sched_entity **)ptr;
        ptr += nr_cpu_ids * sizeof(void **);

        root_task_group.cfs_rq = (struct cfs_rq **)ptr;
        ptr += nr_cpu_ids * sizeof(void **);
#endif
#ifdef CONFIG_RT_GROUP_SCHED
        root_task_group.rt_se = (struct sched_rt_entity **)ptr;
        ptr += nr_cpu_ids * sizeof(void **);

        root_task_group.rt_rq = (struct rt_rq **)ptr;
        ptr += nr_cpu_ids * sizeof(void **);

#endif
```

As I already mentioned, the Linux group scheduling mechanism allows to specify a hierarchy. The root of such hierarchies is the `root_runqueuetask_group` task group structure. This structure contains many fields, but we are interested in `se`, `rt_se`, `cfs_rq` and `rt_rq` for this moment:

The first two are instances of `sched_entity` structure. It is defined in the include/linux/sched.h kernel header filed and used by the scheduler as an unit of scheduling.

```
struct task_group {
    ...
    ...
    struct sched_entity **se;
    struct cfs_rq **cfs_rq;
    ...
    ...
}
```

The `cfs_rq` and `rt_rq` present `run queues`. A `run queue` is a special `per-cpu` structure that is used by the Linux kernel scheduler to store `active` threads or in other words set of threads which potentially will be picked up by the scheduler to run.

The space is allocated and the next step is to initialize a `bandwidth` of CPU for `real-time` and `deadline` tasks:

```
init_rt_bandwidth(&def_rt_bandwidth,
                  global_rt_period(), global_rt_runtime());
init_dl_bandwidth(&def_dl_bandwidth,
                  global_rt_period(), global_rt_runtime());
```

All groups have to be able to rely on the amount of CPU time. The two following structures: `def_rt_bandwidth` and `def_dl_bandwidth` represent default values of bandwidths for `real-time` and `deadline` tasks. We will not look at definition of these structures as it is not so important for now, but we are interested in two following values:

- `sched_rt_period_us`;
- `sched_rt_runtime_us`.

The first represents a period and the second represents quantum that is allocated for `real-time` tasks during `sched_rt_period_us`. You may see global values of these parameters in the:

```
$ cat /proc/sys/kernel/sched_rt_period_us
1000000

$ cat /proc/sys/kernel/sched_rt_runtime_us
950000
```

The values related to a group can be configured in `<cgroup>/cpu.rt_period_us` and `<cgroup>/cpu.rt_runtime_us`. Due no one filesystem is not mounted yet, the `def_rt_bandwidth` and the `def_dl_bandwidth` will be initialzed with default values which will be retuned by the `global_rt_period` and `global_rt_runtime` functions.

That's all with the bandwiths of `real-time` and `deadline` tasks and in the next step, depends on enable of SMP, we make initialization of the `root domain`:

```
#ifdef CONFIG_SMP
    init_defrootdomain();
#endif
```

The real-time scheduler requires global resources to make scheduling decision. But unfortunately scalability bottlenecks appear as the number of CPUs increase. The concept of `root domains` was introduced for improving scalability and avoid such bottlenecks. Instead of bypassing over all `run queues`, the scheduler gets information about a CPU where/from to push/pull a `real-time` task from the `root_domain` structure. This structure is defined in the kernel/sched/sched.h kernel header file and just keeps track of CPUs that can be used to push or pull a process.

After `root domain` initialization, we make initialization of the `bandwidth` for the `real-time` tasks of the `root task group` as we did the same above:

```
#ifdef CONFIG_RT_GROUP_SCHED
    init_rt_bandwidth(&root_task_group.rt_bandwidth,
            global_rt_period(), global_rt_runtime());
#endif
```

with the same default values.

In the next step, depends on the `CONFIG_CGROUP_SCHED` kernel configuration option we allocate `slab` cache for `task_group(s)` and initialize the `siblings` and `children` lists of the root task group. As we can read from the documentation, the `CONFIG_CGROUP_SCHED` is:

```
This option allows you to create arbitrary task groups using the "cgroup" pseudo
filesystem and control the cpu bandwidth allocated to each such task group.
```

As we finished with the lists initialization, we can see the call of the `autogroup_init` function:

```
#ifdef CONFIG_CGROUP_SCHED
        list_add(&root_task_group.list, &task_groups);
        INIT_LIST_HEAD(&root_task_group.children);
        INIT_LIST_HEAD(&root_task_group.siblings);
        autogroup_init(&init_task);
#endif
```

which initializes automatic process group scheduling. The `autogroup` feature is about automatic creation and population of a new task group during creation of a new session via setsid call.

After this we are going through the all `possible` CPUs (you can remember that `possible` CPUs are stored in the `cpu_possible_mask` bitmap that can ever be available in the system) and initialize a `runqueue` for each `possible` cpu:

```
for_each_possible_cpu(i) {
    struct rq *rq;
    ...
    ...
    ...
```

The `rq` structure in the Linux kernel is defined in the kernel/sched/sched.h. As I already mentioned this above, a `run queue` is a fundamental data structure in a scheduling process. The scheduler uses it to determine who will be runned next. As you may see, this structure has many different fields and we will not cover all of them here, but we will look on them when they will be directly used.

After initialization of `per-cpu` run queues with default values, we need to setup `load weight` of the first task in the system:

```
set_load_weight(&init_task);
```

First of all let's try to understand what is it `load weight` of a process. If you will look at the definition of the `sched_entity` structure, you will see that it starts from the `load` field:

```
struct sched_entity {
    struct load_weight        load;
    ...
    ...
```

```
    ...
}
```

represented by the `load_weight` structure which just contains two fields that represent actual load weight of a scheduler entity and its invariant value:

```
struct load_weight {
    unsigned long    weight;
    u32              inv_weight;
};
```

You already may know that each process in the system has `priority`. The higher priority allows to get more time to run. A `load weight` of a process is a relation between priority of this process and timeslices of this process. Each process has three following fields related to priority:

```
struct task_struct {
...
...
...
    int              prio;
    int              static_prio;
    int              normal_prio;
...
...
...
}
```

The first one is `dynamic priority` which can't be changed during lifetime of a process based on its static priority and interactivity of the process. The `static_prio` contains initial priority most likely well-known to you `nice value`. This value does not changed by the kernel if a user will not change it. The last one is `normal_priority` based on the value of the `static_prio` too, but also it depends on the scheduling policy of a process.

So the main goal of the `set_load_weight` function is to initialze `load_weight` fields for the `init` task:

```
static void set_load_weight(struct task_struct *p)
{
    int prio = p->static_prio - MAX_RT_PRIO;
    struct load_weight *load = &p->se.load;

    if (idle_policy(p->policy)) {
        load->weight = scale_load(WEIGHT_IDLEPRIO);
        load->inv_weight = WMULT_IDLEPRIO;
        return;
    }

    load->weight = scale_load(sched_prio_to_weight[prio]);
    load->inv_weight = sched_prio_to_wmult[prio];
}
```

As you may see we calculate initial `prio` from the initial value of the `static_prio` of the `init` task and use it as index of `sched_prio_to_weight` and `sched_prio_to_wmult` arrays to set `weight` and `inv_weight` values. These two arrays contain a `load weight` depends on priority value. In a case of when a process is `idle` process, we set minimal load weight.

For this moment we came to the end of initialization process of the Linux kernel scheduler. The last steps are: to make current process (it will be the first `init` process) `idle` that will be runned when a cpu has no other process to run. Calculating next time period of the next calculation of CPU load and initialization of the `fair` class:

```
__init void init_sched_fair_class(void)
{
```

```
#ifdef CONFIG_SMP
    open_softirq(SCHED_SOFTIRQ, run_rebalance_domains);
#endif
}
```

Here we register a soft irq that will call the `run_rebalance_domains` handler. After the `SCHED_SOFTIRQ` will be triggered, the `run_rebalance` will be called to rebalance a run queue on the current CPU.

The last two steps of the `sched_init` function is to initialization of scheduler statistics and setting `scheeduler_running` variable:

```
scheduler_running = 1;
```

That's all. Linux kernel scheduler is initialized. Of course, we have skipped many different details and explanations here, because we need to know and understand how different concepts (like process and process groups, runqueue, rcu, etc.) works in the linux kernel , but we took a short look on the scheduler initialization process. We will look all other details in the separate part which will be fully dedicated to the scheduler.

Conclusion

It is the end of the eighth part about the linux kernel initialization process. In this part, we looked on the initialization process of the scheduler and we will continue in the next part to dive in the linux kernel initialization process and will see initialization of the RCU and many other initialization stuff in the next part.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- CPU masks
- high-resolution kernel timer
- spinlock
- Run queue
- Linux kernel memory manager
- slub
- virtual file system
- Linux kernel hotplug documentation
- IRQ
- Global Descriptor Table
- Per-CPU variables
- SMP
- RCU
- CFS Scheduler documentation
- Real-Time group scheduling
- Previous part

Kernel initialization. Part 9.

RCU initialization

This is ninth part of the Linux Kernel initialization process and in the previous part we stopped at the scheduler initialization. In this part we will continue to dive to the linux kernel initialization process and the main purpose of this part will be to learn about initialization of the RCU. We can see that the next step in the init/main.c after the `sched_init` is the call of the `preempt_disable`. There are two macros:

- `preempt_disable`
- `preempt_enable`

for preemption disabling and enabling. First of all let's try to understand what is `preempt` in the context of an operating system kernel. In simple words, preemption is ability of the operating system kernel to preempt current task to run task with higher priority. Here we need to disable preemption because we will have only one `init` process for the early boot time and we don't need to stop it before we call `cpu_idle` function. The `preempt_disable` macro is defined in the include/linux/preempt.h and depends on the `CONFIG_PREEMPT_COUNT` kernel configuration option. This macro is implemented as:

```
#define preempt_disable() \
do { \
        preempt_count_inc(); \
        barrier(); \
} while (0)
```

and if `CONFIG_PREEMPT_COUNT` is not set just:

```
#define preempt_disable()                       barrier()
```

Let's look on it. First of all we can see one difference between these macro implementations. The `preempt_disable` with `CONFIG_PREEMPT_COUNT` set contains the call of the `preempt_count_inc`. There is special `percpu` variable which stores the number of held locks and `preempt_disable` calls:

```
DECLARE_PER_CPU(int, __preempt_count);
```

In the first implementation of the `preempt_disable` we increment this `__preempt_count`. There is API for returning value of the `__preempt_count`, it is the `preempt_count` function. As we called `preempt_disable`, first of all we increment preemption counter with the `preempt_count_inc` macro which expands to the:

```
#define preempt_count_inc() preempt_count_add(1)
#define preempt_count_add(val)  __preempt_count_add(val)
```

where `preempt_count_add` calls the `raw_cpu_add_4` macro which adds `1` to the given `percpu` variable (`__preempt_count`) in our case (more about `precpu` variables you can read in the part about Per-CPU variables). Ok, we increased `__preempt_count` and the next step we can see the call of the `barrier` macro in the both macros. The `barrier` macro inserts an optimization barrier. In the processors with `x86_64` architecture independent memory access operations can be performed in any order. That's why we need the opportunity to point compiler and processor on compliance of order. This mechanism is memory barrier. Let's consider a simple example:

```
preempt_disable();
foo();
```

```
preempt_enable();
```

Compiler can rearrange it as:

```
preempt_disable();
preempt_enable();
foo();
```

In this case non-preemptible function `foo` can be preempted. As we put `barrier` macro in the `preempt_disable` and `preempt_enable` macros, it prevents the compiler from swapping `preempt_count_inc` with other statements. More about barriers you can read here and here.

In the next step we can see following statement:

```
if (WARN(!irqs_disabled(),
     "Interrupts were enabled *very* early, fixing it\n"))
    local_irq_disable();
```

which check IRQs state, and disabling (with `cli` instruction for `x86_64`) if they are enabled.

That's all. Preemption is disabled and we can go ahead.

Initialization of the integer ID management

In the next step we can see the call of the `idr_init_cache` function which defined in the lib/idr.c. The `idr` library is used in a various places in the linux kernel to manage assigning integer `IDs` to objects and looking up objects by id.

Let's look on the implementation of the `idr_init_cache` function:

```
void __init idr_init_cache(void)
{
        idr_layer_cache = kmem_cache_create("idr_layer_cache",
                                sizeof(struct idr_layer), 0, SLAB_PANIC, NULL);
}
```

Here we can see the call of the `kmem_cache_create` . We already called the `kmem_cache_init` in the init/main.c. This function create generalized caches again using the `kmem_cache_alloc` (more about caches we will see in the Linux kernel memory management chapter). In our case, as we are using `kmem_cache_t` which will be used by the slab allocator and `kmem_cache_create` creates it. As you can see we pass five parameters to the `kmem_cache_create` :

- name of the cache;
- size of the object to store in cache;
- offset of the first object in the page;
- flags;
- constructor for the objects.

and it will create `kmem_cache` for the integer IDs. Integer `IDs` is commonly used pattern to map set of integer IDs to the set of pointers. We can see usage of the integer IDs in the i2c drivers subsystem. For example drivers/i2c/i2c-core.c which represents the core of the `i2c` subsystem defines `ID` for the `i2c` adapter with the `DEFINE_IDR` macro:

```
static DEFINE_IDR(i2c_adapter_idr);
```

and then uses it for the declaration of the `i2c` adapter:

```
static int __i2c_add_numbered_adapter(struct i2c_adapter *adap)
{
  int     id;
  ...
  ...
  ...
  id = idr_alloc(&i2c_adapter_idr, adap, adap->nr, adap->nr + 1, GFP_KERNEL);
  ...
  ...
  ...
}
```

and `id2_adapter_idr` presents dynamically calculated bus number.

More about integer ID management you can read here.

RCU initialization

The next step is RCU initialization with the `rcu_init` function and it's implementation depends on two kernel configuration options:

- `CONFIG_TINY_RCU`
- `CONFIG_TREE_RCU`

In the first case `rcu_init` will be in the kernel/rcu/tiny.c and in the second case it will be defined in the kernel/rcu/tree.c. We will see the implementation of the `tree rcu`, but first of all about the `RCU` in general.

`RCU` or read-copy update is a scalable high-performance synchronization mechanism implemented in the Linux kernel. On the early stage the linux kernel provided support and environment for the concurrently running applications, but all execution was serialized in the kernel using a single global lock. In our days linux kernel has no single global lock, but provides different mechanisms including lock-free data structures, percpu data structures and other. One of these mechanisms is - the `read-copy update`. The `RCU` technique is designed for rarely-modified data structures. The idea of the `RCU` is simple. For example we have a rarely-modified data structure. If somebody wants to change this data structure, we make a copy of this data structure and make all changes in the copy. In the same time all other users of the data structure use old version of it. Next, we need to choose safe moment when original version of the data structure will have no users and update it with the modified copy.

Of course this description of the `RCU` is very simplified. To understand some details about `RCU`, first of all we need to learn some terminology. Data readers in the `RCU` executed in the critical section. Every time when data reader get to the critical section, it calls the `rcu_read_lock`, and `rcu_read_unlock` on exit from the critical section. If the thread is not in the critical section, it will be in state which called - `quiescent state`. The moment when every thread is in the `quiescent state` called - `grace period`. If a thread wants to remove an element from the data structure, this occurs in two steps. First step is `removal` - atomically removes element from the data structure, but does not release the physical memory. After this thread-writer announces and waits until it is finished. From this moment, the removed element is available to the thread-readers. After the `grace period` finished, the second step of the element removal will be started, it just removes the element from the physical memory.

There a couple of implementations of the `RCU`. Old `RCU` called classic, the new implementation called `tree` RCU. As you may already understand, the `CONFIG_TREE_RCU` kernel configuration option enables tree `RCU`. Another is the `tiny` RCU which depends on `CONFIG_TINY_RCU` and `CONFIG_SMP=n`. We will see more details about the `RCU` in general in the separate chapter about synchronization primitives, but now let's look on the `rcu_init` implementation from the kernel/rcu/tree.c:

```
void __init rcu_init(void)
{
         int cpu;

         rcu_bootup_announce();
         rcu_init_geometry();
         rcu_init_one(&rcu_bh_state, &rcu_bh_data);
```

```
        rcu_init_one(&rcu_sched_state, &rcu_sched_data);
        __rcu_init_preempt();
        open_softirq(RCU_SOFTIRQ, rcu_process_callbacks);

        /*
         * We don't need protection against CPU-hotplug here because
         * this is called early in boot, before either interrupts
         * or the scheduler are operational.
         */
        cpu_notifier(rcu_cpu_notify, 0);
        pm_notifier(rcu_pm_notify, 0);
        for_each_online_cpu(cpu)
                rcu_cpu_notify(NULL, CPU_UP_PREPARE, (void *)(long)cpu);

        rcu_early_boot_tests();
}
```

In the beginning of the `rcu_init` function we define `cpu` variable and call `rcu_bootup_announce`. The `rcu_bootup_announce` function is pretty simple:

```
static void __init rcu_bootup_announce(void)
{
        pr_info("Hierarchical RCU implementation.\n");
        rcu_bootup_announce_oddness();
}
```

It just prints information about the `RCU` with the `pr_info` function and `rcu_bootup_announce_oddness` which uses `pr_info` too, for printing different information about the current `RCU` configuration which depends on different kernel configuration options like `CONFIG_RCU_TRACE`, `CONFIG_PROVE_RCU`, `CONFIG_RCU_FANOUT_EXACT`, etc. In the next step, we can see the call of the `rcu_init_geometry` function. This function is defined in the same source code file and computes the node tree geometry depends on the amount of CPUs. Actually `RCU` provides scalability with extremely low internal RCU lock contention. What if a data structure will be read from the different CPUs? `RCU` API provides the `rcu_state` structure which presents RCU global state including node hierarchy. Hierarchy is presented by the:

```
struct rcu_node node[NUM_RCU_NODES];
```

array of structures. As we can read in the comment of above definition:

```
The root (first level) of the hierarchy is in ->node[0] (referenced by ->level[0]), the second
level in ->node[1] through ->node[m] (->node[1] referenced by ->level[1]), and the third level
in ->node[m+1] and following (->node[m+1] referenced by ->level[2]).  The number of levels is
determined by the number of CPUs and by CONFIG_RCU_FANOUT.

Small systems will have a "hierarchy" consisting of a single rcu_node.
```

The `rcu_node` structure is defined in the kernel/rcu/tree.h and contains information about current grace period, is grace period completed or not, CPUs or groups that need to switch in order for current grace period to proceed, etc. Every `rcu_node` contains a lock for a couple of CPUs. These `rcu_node` structures are embedded into a linear array in the `rcu_state` structure and represented as a tree with the root as the first element and covers all CPUs. As you can see the number of the rcu nodes determined by the `NUM_RCU_NODES` which depends on number of available CPUs:

```
#define NUM_RCU_NODES (RCU_SUM - NR_CPUS)
#define RCU_SUM (NUM_RCU_LVL_0 + NUM_RCU_LVL_1 + NUM_RCU_LVL_2 + NUM_RCU_LVL_3 + NUM_RCU_LVL_4)
```

where levels values depend on the `CONFIG_RCU_FANOUT_LEAF` configuration option. For example for the simplest case, one `rcu_node` will cover two CPU on machine with the eight CPUs:

```
+-----------------------------------------------------------------+
|  rcu_state                                                      |
|                 +----------------------+                        |
|                 |         root         |                        |
|                 |       rcu_node       |                        |
|                 +----------------------+                        |
|                    |                |                          |
|               +----v-----+       +--v-------+                  |
|               |          |       |          |                  |
|               | rcu_node |       | rcu_node |                  |
|               |          |       |          |                  |
|         +------------------+     +----------------+            |
|         |                  |        |             |            |
|         |                  |        |             |            |
|    +----v-----+    +-------v--+   +-v--------+  +-v--------+    |
|    |          |    |          |   |          |  |          |    |
|    | rcu_node |    | rcu_node |   | rcu_node |  | rcu_node |    |
|    |          |    |          |   |          |  |          |    |
|    +----------+    +----------+   +----------+  +----------+    |
|         |                 |             |               |      |
|         |                 |             |               |      |
|         |                 |             |               |      |
|         |                 |             |               |      |
+---------|-----------------|-------------|---------------|-------+
          |                 |             |               |
+---------v-----------------v-------------v---------------v--------+
|                 |                |               |               |
|     CPU1        |      CPU3      |      CPU5     |     CPU7      |
|                 |                |               |               |
|     CPU2        |      CPU4      |      CPU6     |     CPU8      |
|                 |                |               |               |
+------------------------------------------------------------------+
```

So, in the `rcu_init_geometry` function we just need to calculate the total number of `rcu_node` structures. We start to do it with the calculation of the `jiffies` till to the first and next `fqs` which is `force-quiescent-state` (read above about it):

```
d = RCU_JIFFIES_TILL_FORCE_QS + nr_cpu_ids / RCU_JIFFIES_FQS_DIV;
if (jiffies_till_first_fqs == ULONG_MAX)
        jiffies_till_first_fqs = d;
if (jiffies_till_next_fqs == ULONG_MAX)
        jiffies_till_next_fqs = d;
```

where:

```
#define RCU_JIFFIES_TILL_FORCE_QS (1 + (HZ > 250) + (HZ > 500))
#define RCU_JIFFIES_FQS_DIV     256
```

As we calculated these jiffies, we check that previous defined `jiffies_till_first_fqs` and `jiffies_till_next_fqs` variables are equal to the ULONG_MAX (their default values) and set they equal to the calculated value. As we did not touch these variables before, they are equal to the `ULONG_MAX`:

```
static ulong jiffies_till_first_fqs = ULONG_MAX;
static ulong jiffies_till_next_fqs = ULONG_MAX;
```

In the next step of the `rcu_init_geometry`, we check that `rcu_fanout_leaf` didn't change (it has the same value as `CONFIG_RCU_FANOUT_LEAF` in compile-time) and equal to the value of the `CONFIG_RCU_FANOUT_LEAF` configuration option, we just return:

```
if (rcu_fanout_leaf == CONFIG_RCU_FANOUT_LEAF &&
    nr_cpu_ids == NR_CPUS)
```

```
    return;
```

After this we need to compute the number of nodes that an `rcu_node` tree can handle with the given number of levels:

```
rcu_capacity[0] = 1;
rcu_capacity[1] = rcu_fanout_leaf;
for (i = 2; i <= MAX_RCU_LVLS; i++)
    rcu_capacity[i] = rcu_capacity[i - 1] * CONFIG_RCU_FANOUT;
```

And in the last step we calculate the number of rcu_nodes at each level of the tree in the loop.

As we calculated geometry of the `rcu_node` tree, we need to go back to the `rcu_init` function and next step we need to initialize two `rcu_state` structures with the `rcu_init_one` function:

```
rcu_init_one(&rcu_bh_state, &rcu_bh_data);
rcu_init_one(&rcu_sched_state, &rcu_sched_data);
```

The `rcu_init_one` function takes two arguments:

- Global `RCU` state;
- Per-CPU data for `RCU`.

Both variables defined in the kernel/rcu/tree.h with its `percpu` data:

```
extern struct rcu_state rcu_bh_state;
DECLARE_PER_CPU(struct rcu_data, rcu_bh_data);
```

About this states you can read here. As I wrote above we need to initialize `rcu_state` structures and `rcu_init_one` function will help us with it. After the `rcu_state` initialization, we can see the call of the `__rcu_init_preempt` which depends on the `CONFIG_PREEMPT_RCU` kernel configuration option. It does the same as previous functions - initialization of the `rcu_preempt_state` structure with the `rcu_init_one` function which has `rcu_state` type. After this, in the `rcu_init`, we can see the call of the:

```
open_softirq(RCU_SOFTIRQ, rcu_process_callbacks);
```

function. This function registers a handler of the `pending interrupt`. Pending interrupt or `softirq` supposes that part of actions can be delayed for later execution when the system is less loaded. Pending interrupts is represented by the following structure:

```
struct softirq_action
{
        void    (*action)(struct softirq_action *);
};
```

which is defined in the include/linux/interrupt.h and contains only one field - handler of an interrupt. You can check about `softirqs` in the your system with the:

```
$ cat /proc/softirqs
                    CPU0       CPU1       CPU2       CPU3       CPU4       CPU5       CPU6       CPU7
          HI:          2          0          0          1          0          2          0          0
       TIMER:     137779     108110     139573     107647     107408     114972      99653      98665
      NET_TX:       1127          0          4          0          1          1          0          0
      NET_RX:        334        221     132939       3076        451        361        292        303
       BLOCK:       5253       5596          8        779       2016      37442         28       2855
BLOCK_IOPOLL:          0          0          0          0          0          0          0          0
     TASKLET:         66          0       2916        113          0         24      26708          0
```

```
        SCHED:     102350      75950      91705      75356      75323      82627      69279      69914
      HRTIMER:        510        302        368        260        219        255        248        246
          RCU:      81290      68062      82979      69015      68390      69385      63304      63473
```

The `open_softirq` function takes two parameters:

- index of the interrupt;
- interrupt handler.

and adds interrupt handler to the array of the pending interrupts:

```
void open_softirq(int nr, void (*action)(struct softirq_action *))
{
        softirq_vec[nr].action = action;
}
```

In our case the interrupt handler is - `rcu_process_callbacks` which is defined in the kernel/rcu/tree.c and does the `RCU` core processing for the current CPU. After we registered `softirq` interrupt for the `RCU`, we can see the following code:

```
cpu_notifier(rcu_cpu_notify, 0);
pm_notifier(rcu_pm_notify, 0);
for_each_online_cpu(cpu)
    rcu_cpu_notify(NULL, CPU_UP_PREPARE, (void *)(long)cpu);
```

Here we can see registration of the `cpu` notifier which needs in systems which supports CPU hotplug and we will not dive into details about this theme. The last function in the `rcu_init` is the `rcu_early_boot_tests`:

```
void rcu_early_boot_tests(void)
{
        pr_info("Running RCU self tests\n");

        if (rcu_self_test)
                 early_boot_test_call_rcu();
         if (rcu_self_test_bh)
                 early_boot_test_call_rcu_bh();
         if (rcu_self_test_sched)
                early_boot_test_call_rcu_sched();
}
```

which runs self tests for the `RCU`.

That's all. We saw initialization process of the `RCU` subsystem. As I wrote above, more about the `RCU` will be in the separate chapter about synchronization primitives.

Rest of the initialization process

Ok, we already passed the main theme of this part which is `RCU` initialization, but it is not the end of the linux kernel initialization process. In the last paragraph of this theme we will see a couple of functions which work in the initialization time, but we will not dive into deep details around this function for different reasons. Some reasons not to dive into details are following:

- They are not very important for the generic kernel initialization process and depend on the different kernel configuration;
- They have the character of debugging and not important for now;
- We will see many of this stuff in the separate parts/chapters.

After we initialized `RCU`, the next step which you can see in the init/main.c is the - `trace_init` function. As you can understand from its name, this function initialize tracing subsystem. You can read more about linux kernel trace system - here.

After the `trace_init`, we can see the call of the `radix_tree_init`. If you are familiar with the different data structures, you can understand from the name of this function that it initializes kernel implementation of the Radix tree. This function is defined in the lib/radix-tree.c and you can read more about it in the part about Radix tree.

In the next step we can see the functions which are related to the `interrupts handling` subsystem, they are:

* `early_irq_init`
* `init_IRQ`
* `softirq_init`

We will see explanation about this functions and their implementation in the special part about interrupts and exceptions handling. After this many different functions (like `init_timers`, `hrtimers_init`, `time_init`, etc.) which are related to different timing and timers stuff. We will see more about these function in the chapter about timers.

The next couple of functions are related with the perf events - `perf_event-init` (there will be separate chapter about perf), initialization of the `profiling` with the `profile_init`. After this we enable `irq` with the call of the:

```
local_irq_enable();
```

which expands to the `sti` instruction and making post initialization of the SLAB with the call of the `kmem_cache_init_late` function (As I wrote above we will know about the `SLAB` in the Linux memory management chapter).

After the post initialization of the `SLAB`, next point is initialization of the console with the `console_init` function from the drivers/tty/tty_io.c.

After the console initialization, we can see the `lockdep_info` function which prints information about the Lock dependency validator. After this, we can see the initialization of the dynamic allocation of the `debug objects` with the `debug_objects_mem_init`, kernel memory leak detector initialization with the `kmemleak_init`, `percpu` pageset setup with the `setup_per_cpu_pageset`, setup of the NUMA policy with the `numa_policy_init`, setting time for the scheduler with the `sched_clock_init`, `pidmap` initialization with the call of the `pidmap_init` function for the initial `PID` namespace, cache creation with the `anon_vma_init` for the private virtual memory areas and early initialization of the ACPI with the `acpi_early_init`.

This is the end of the ninth part of the linux kernel initialization process and here we saw initialization of the RCU. In the last paragraph of this part (`Rest of the initialization process`) we will go through many functions but did not dive into details about their implementations. Do not worry if you do not know anything about these stuff or you know and do not understand anything about this. As I already wrote many times, we will see details of implementations in other parts or other chapters.

Conclusion

It is the end of the ninth part about the linux kernel initialization process. In this part, we looked on the initialization process of the `RCU` subsystem. In the next part we will continue to dive into linux kernel initialization process and I hope that we will finish with the `start_kernel` function and will go to the `rest_init` function from the same init/main.c source code file and will see the start of the first process.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

* lock-free data structures
* kmemleak

- ACPI
- IRQs
- RCU
- RCU documentation
- integer ID management
- Documentation/memory-barriers.txt
- Runtime locking correctness validator
- Per-CPU variables
- Linux kernel memory management
- slab
- i2c
- Previous part

Kernel initialization. Part 10.

End of the linux kernel initialization process

This is tenth part of the chapter about linux kernel initialization process and in the previous part we saw the initialization of the RCU and stopped on the call of the `acpi_early_init` function. This part will be the last part of the Kernel initialization process chapter, so let's finish it.

After the call of the `acpi_early_init` function from the init/main.c, we can see the following code:

```
#ifdef CONFIG_X86_ESPFIX64
    init_espfix_bsp();
#endif
```

Here we can see the call of the `init_espfix_bsp` function which depends on the `CONFIG_X86_ESPFIX64` kernel configuration option. As we can understand from the function name, it does something with the stack. This function is defined in the arch/x86/kernel/espfix_64.c and prevents leaking of `31:16` bits of the `esp` register during returning to 16-bit stack. First of all we install `espfix` page upper directory into the kernel page directory in the `init_espfix_bs` :

```
pgd_p = &init_level4_pgt[pgd_index(ESPFIX_BASE_ADDR)];
pgd_populate(&init_mm, pgd_p, (pud_t *)espfix_pud_page);
```

Where `ESPFIX_BASE_ADDR` is:

```
#define PGDIR_SHIFT     39
#define ESPFIX_PGD_ENTRY _AC(-2, UL)
#define ESPFIX_BASE_ADDR (ESPFIX_PGD_ENTRY << PGDIR_SHIFT)
```

Also we can find it in the Documentation/x86/x86_64/mm:

```
... unused hole ...
ffffff0000000000 - ffffff7fffffffff (=39 bits) %esp fixup stacks
... unused hole ...
```

After we've filled page global directory with the `espfix` pud, the next step is call of the `init_espfix_random` and `init_espfix_ap` functions. The first function returns random locations for the `espfix` page and the second enables the `espfix` for the current CPU. After the `init_espfix_bsp` finished the work, we can see the call of the `thread_info_cache_init` function which defined in the kernel/fork.c and allocates cache for the `thread_info` if `THREAD_SIZE` is less than `PAGE_SIZE` :

```
# if THREAD_SIZE >= PAGE_SIZE
...
...
...
void thread_info_cache_init(void)
{
        thread_info_cache = kmem_cache_create("thread_info", THREAD_SIZE,
                                              THREAD_SIZE, 0, NULL);
        BUG_ON(thread_info_cache == NULL);
}
...
...
...
```

```
#endif
```

As we already know the `PAGE_SIZE` is `(_AC(1,UL) << PAGE_SHIFT)` or `4096` bytes and `THREAD_SIZE` is `(PAGE_SIZE << THREAD_SIZE_ORDER)` or `16384` bytes for the `x86_64`. The next function after the `thread_info_cache_init` is the `cred_init` from the kernel/cred.c. This function just allocates cache for the credentials (like `uid`, `gid`, etc.):

```
void __init cred_init(void)
{
         cred_jar = kmem_cache_create("cred_jar", sizeof(struct cred),
                                     0, SLAB_HWCACHE_ALIGN|SLAB_PANIC, NULL);
}
```

more about credentials you can read in the Documentation/security/credentials.txt. Next step is the `fork_init` function from the kernel/fork.c. The `fork_init` function allocates cache for the `task_struct`. Let's look on the implementation of the `fork_init`. First of all we can see definitions of the `ARCH_MIN_TASKALIGN` macro and creation of a slab where task_structs will be allocated:

```
#ifndef CONFIG_ARCH_TASK_STRUCT_ALLOCATOR
#ifndef ARCH_MIN_TASKALIGN
#define ARCH_MIN_TASKALIGN      L1_CACHE_BYTES
#endif
        task_struct_cachep =
                kmem_cache_create("task_struct", sizeof(struct task_struct),
                        ARCH_MIN_TASKALIGN, SLAB_PANIC | SLAB_NOTRACK, NULL);
#endif
```

As we can see this code depends on the `CONFIG_ARCH_TASK_STRUCT_ACLLOCATOR` kernel configuration option. This configuration option shows the presence of the `alloc_task_struct` for the given architecture. As `x86_64` has no `alloc_task_struct` function, this code will not work and even will not be compiled on the `x86_64`.

Allocating cache for init task

After this we can see the call of the `arch_task_cache_init` function in the `fork_init`:

```
void arch_task_cache_init(void)
{
        task_xstate_cachep =
                kmem_cache_create("task_xstate", xstate_size,
                                  __alignof__(union thread_xstate),
                                  SLAB_PANIC | SLAB_NOTRACK, NULL);
        setup_xstate_comp();
}
```

The `arch_task_cache_init` does initialization of the architecture-specific caches. In our case it is `x86_64`, so as we can see, the `arch_task_cache_init` allocates cache for the `task_xstate` which represents FPU state and sets up offsets and sizes of all extended states in xsave area with the call of the `setup_xstate_comp` function. After the `arch_task_cache_init` we calculate default maximum number of threads with the:

```
set_max_threads(MAX_THREADS);
```

where default maximum number of threads is:

```
#define FUTEX_TID_MASK  0x3fffffff
#define MAX_THREADS     FUTEX_TID_MASK
```

In the end of the `fork_init` function we initialize signal handler:

```
init_task.signal->rlim[RLIMIT_NPROC].rlim_cur = max_threads/2;
init_task.signal->rlim[RLIMIT_NPROC].rlim_max = max_threads/2;
init_task.signal->rlim[RLIMIT_SIGPENDING] =
        init_task.signal->rlim[RLIMIT_NPROC];
```

As we know the `init_task` is an instance of the `task_struct` structure, so it contains `signal` field which represents signal handler. It has following type `struct signal_struct`. On the first two lines we can see setting of the current and maximum limit of the `resource limits`. Every process has an associated set of resource limits. These limits specify amount of resources which current process can use. Here `rlim` is resource control limit and presented by the:

```
struct rlimit {
        __kernel_ulong_t        rlim_cur;
        __kernel_ulong_t        rlim_max;
};
```

structure from the include/uapi/linux/resource.h. In our case the resource is the `RLIMIT_NPROC` which is the maximum number of processes that user can own and `RLIMIT_SIGPENDING` - the maximum number of pending signals. We can see it in the:

```
cat /proc/self/limits
Limit                     Soft Limit           Hard Limit           Units
...
...
...
Max processes             63815                63815                processes
Max pending signals       63815                63815                signals
...
...
...
```

Initialization of the caches

The next function after the `fork_init` is the `proc_caches_init` from the kernel/fork.c. This function allocates caches for the memory descriptors (or `mm_struct` structure). At the beginning of the `proc_caches_init` we can see allocation of the different SLAB caches with the call of the `kmem_cache_create`:

- `sighand_cachep` - manage information about installed signal handlers;
- `signal_cachep` - manage information about process signal descriptor;
- `files_cachep` - manage information about opened files;
- `fs_cachep` - manage filesystem information.

After this we allocate `SLAB` cache for the `mm_struct` structures:

```
mm_cachep = kmem_cache_create("mm_struct",
                         sizeof(struct mm_struct), ARCH_MIN_MMSTRUCT_ALIGN,
                         SLAB_HWCACHE_ALIGN|SLAB_PANIC|SLAB_NOTRACK, NULL);
```

After this we allocate `SLAB` cache for the important `vm_area_struct` which used by the kernel to manage virtual memory space:

```
vm_area_cachep = KMEM_CACHE(vm_area_struct, SLAB_PANIC);
```

Note, that we use `KMEM_CACHE` macro here instead of the `kmem_cache_create`. This macro is defined in the include/linux/slab.h and just expands to the `kmem_cache_create` call:

```
#define KMEM_CACHE(__struct, __flags) kmem_cache_create(#__struct,\
                sizeof(struct __struct), __alignof__(struct __struct),\
                (__flags), NULL)
```

The `KMEM_CACHE` has one difference from `kmem_cache_create`. Take a look on `__alignof__` operator. The `KMEM_CACHE` macro aligns `SLAB` to the size of the given structure, but `kmem_cache_create` uses given value to align space. After this we can see the call of the `mmap_init` and `nsproxy_cache_init` functions. The first function initializes virtual memory area `SLAB` and the second function initializes `SLAB` for namespaces.

The next function after the `proc_caches_init` is `buffer_init`. This function is defined in the fs/buffer.c source code file and allocate cache for the `buffer_head`. The `buffer_head` is a special structure which defined in the include/linux/buffer_head.h and used for managing buffers. In the start of the `buffer_init` function we allocate cache for the `struct buffer_head` structures with the call of the `kmem_cache_create` function as we did in the previous functions. And calculate the maximum size of the buffers in memory with:

```
nrpages = (nr_free_buffer_pages() * 10) / 100;
max_buffer_heads = nrpages * (PAGE_SIZE / sizeof(struct buffer_head));
```

which will be equal to the `10%` of the `ZONE_NORMAL` (all RAM from the 4GB on the `x86_64`). The next function after the `buffer_init` is - `vfs_caches_init`. This function allocates `SLAB` caches and hashtable for different VFS caches. We already saw the `vfs_caches_init_early` function in the eighth part of the linux kernel initialization process which initialized caches for `dcache` (or directory-cache) and inode cache. The `vfs_caches_init` function makes post-early initialization of the `dcache` and `inode` caches, private data cache, hash tables for the mount points, etc. More details about VFS will be described in the separate part. After this we can see `signals_init` function. This function is defined in the kernel/signal.c and allocates a cache for the `sigqueue` structures which represents queue of the real time signals. The next function is `page_writeback_init`. This function initializes the ratio for the dirty pages. Every low-level page entry contains the `dirty` bit which indicates whether a page has been written to after been loaded into memory.

Creation of the root for the procfs

After all of this preparations we need to create the root for the proc filesystem. We will do it with the call of the `proc_root_init` function from the fs/proc/root.c. At the start of the `proc_root_init` function we allocate the cache for the inodes and register a new filesystem in the system with the:

```
err = register_filesystem(&proc_fs_type);
      if (err)
                return;
```

As I wrote above we will not dive into details about VFS and different filesystems in this chapter, but will see it in the chapter about the `VFS`. After we've registered a new filesystem in our system, we call the `proc_self_init` function from the fs/proc/self.c and this function allocates `inode` number for the `self` (`/proc/self` directory refers to the process accessing the `/proc` filesystem). The next step after the `proc_self_init` is `proc_setup_thread_self` which setups the `/proc/thread-self` directory which contains information about current thread. After this we create `/proc/self/mounts` symlink which will contains mount points with the call of the

```
proc_symlink("mounts", NULL, "self/mounts");
```

and a couple of directories depends on the different configuration options:

```
#ifdef CONFIG_SYSVIPC
        proc_mkdir("sysvipc", NULL);
#endif
```

```
        proc_mkdir("fs", NULL);
        proc_mkdir("driver", NULL);
        proc_mkdir("fs/nfsd", NULL);
#if defined(CONFIG_SUN_OPENPROMFS) || defined(CONFIG_SUN_OPENPROMFS_MODULE)
        proc_mkdir("openprom", NULL);
#endif
        proc_mkdir("bus", NULL);
        ...
        ...
        ...
        if (!proc_mkdir("tty", NULL))
                return;
        proc_mkdir("tty/ldisc", NULL);
        ...
        ...
        ...
```

In the end of the `proc_root_init` we call the `proc_sys_init` function which creates `/proc/sys` directory and initializes the Sysctl.

It is the end of `start_kernel` function. I did not describe all functions which are called in the `start_kernel`. I skipped them, because they are not important for the generic kernel initialization stuff and depend on only different kernel configurations. They are `taskstats_init_early` which exports per-task statistic to the user-space, `delayacct_init` - initializes per-task delay accounting, `key_init` and `security_init` initialize different security stuff, `check_bugs` - fix some architecture-dependent bugs, `ftrace_init` function executes initialization of the ftrace, `cgroup_init` makes initialization of the rest of the cgroup subsystem,etc. Many of these parts and subsystems will be described in the other chapters.

That's all. Finally we have passed through the long-long `start_kernel` function. But it is not the end of the linux kernel initialization process. We haven't run the first process yet. In the end of the `start_kernel` we can see the last call of the - `rest_init` function. Let's go ahead.

First steps after the start_kernel

The `rest_init` function is defined in the same source code file as `start_kernel` function, and this file is init/main.c. In the beginning of the `rest_init` we can see call of the two following functions:

```
    rcu_scheduler_starting();
    smpboot_thread_init();
```

The first `rcu_scheduler_starting` makes RCU scheduler active and the second `smpboot_thread_init` registers the `smpboot_thread_notifier` CPU notifier (more about it you can read in the CPU hotplug documentation. After this we can see the following calls:

```
kernel_thread(kernel_init, NULL, CLONE_FS);
pid = kernel_thread(kthreadd, NULL, CLONE_FS | CLONE_FILES);
```

Here the `kernel_thread` function (defined in the kernel/fork.c) creates new kernel thread.As we can see the `kernel_thread` function takes three arguments:

- Function which will be executed in a new thread;
- Parameter for the `kernel_init` function;
- Flags.

We will not dive into details about `kernel_thread` implementation (we will see it in the chapter which describe scheduler, just need to say that `kernel_thread` invokes clone). Now we only need to know that we create new kernel thread with `kernel_thread` function, parent and child of the thread will use shared information about filesystem and it will start to execute `kernel_init` function. A kernel thread differs from a user thread that it runs in kernel mode. So with these two `kernel_thread` calls we create two new kernel threads with the `PID = 1` for `init` process and `PID = 2` for `kthreadd`. We already know what is `init` process. Let's look on the `kthreadd`. It is a special kernel thread which manages and helps different parts of the kernel to create another kernel thread. We can see it in the output of the `ps` util:

```
$ ps -ef | grep kthreadd
root         2     0  0 Jan11 ?        00:00:00 [kthreadd]
```

Let's postpone `kernel_init` and `kthreadd` for now and go ahead in the `rest_init`. In the next step after we have created two new kernel threads we can see the following code:

```
    rcu_read_lock();
    kthreadd_task = find_task_by_pid_ns(pid, &init_pid_ns);
    rcu_read_unlock();
```

The first `rcu_read_lock` function marks the beginning of an RCU read-side critical section and the `rcu_read_unlock` marks the end of an RCU read-side critical section. We call these functions because we need to protect the `find_task_by_pid_ns`. The `find_task_by_pid_ns` returns pointer to the `task_struct` by the given pid. So, here we are getting the pointer to the `task_struct` for `PID = 2` (we got it after `kthreadd` creation with the `kernel_thread`). In the next step we call `complete` function

```
complete(&kthreadd_done);
```

and pass address of the `kthreadd_done`. The `kthreadd_done` defined as

```
static __initdata DECLARE_COMPLETION(kthreadd_done);
```

where `DECLARE_COMPLETION` macro defined as:

```
#define DECLARE_COMPLETION(work) \
         struct completion work = COMPLETION_INITIALIZER(work)
```

and expands to the definition of the `completion` structure. This structure is defined in the include/linux/completion.h and presents `completions` concept. Completions is a code synchronization mechanism which provides race-free solution for the threads that must wait for some process to have reached a point or a specific state. Using completions consists of three parts: The first is definition of the `complete` structure and we did it with the `DECLARE_COMPLETION`. The second is call of the `wait_for_completion`. After the call of this function, a thread which called it will not continue to execute and will wait while other thread did not call `complete` function. Note that we call `wait_for_completion` with the `kthreadd_done` in the beginning of the `kernel_init_freeable`:

```
wait_for_completion(&kthreadd_done);
```

And the last step is to call `complete` function as we saw it above. After this the `kernel_init_freeable` function will not be executed while `kthreadd` thread will not be set. After the `kthreadd` was set, we can see three following functions in the `rest_init`:

```
    init_idle_bootup_task(current);
    schedule_preempt_disabled();
    cpu_startup_entry(CPUHP_ONLINE);
```

The first `init_idle_bootup_task` function from the kernel/sched/core.c sets the Scheduling class for the current process (`idle` class in our case):

```
void init_idle_bootup_task(struct task_struct *idle)
{
         idle->sched_class = &idle_sched_class;
}
```

where `idle` class is a low task priority and tasks can be run only when the processor doesn't have anything to run besides this tasks. The second function `schedule_preempt_disabled` disables preempt in `idle` tasks. And the third function `cpu_startup_entry` is defined in the kernel/sched/idle.c and calls `cpu_idle_loop` from the kernel/sched/idle.c. The `cpu_idle_loop` function works as process with `PID = 0` and works in the background. Main purpose of the `cpu_idle_loop` is to consume the idle CPU cycles. When there is no process to run, this process starts to work. We have one process with `idle` scheduling class (we just set the `current` task to the `idle` with the call of the `init_idle_bootup_task` function), so the `idle` thread does not do useful work but just checks if there is an active task to switch to:

```
static void cpu_idle_loop(void)
{
        ...
        ...
        ...
        while (1) {
                while (!need_resched()) {
                ...
                ...
                ...
                }
        ...
        }
```

More about it will be in the chapter about scheduler. So for this moment the `start_kernel` calls the `rest_init` function which spawns an `init` (`kernel_init` function) process and become `idle` process itself. Now is time to look on the `kernel_init` . Execution of the `kernel_init` function starts from the call of the `kernel_init_freeable` function. The `kernel_init_freeable` function first of all waits for the completion of the `kthreadd` setup. I already wrote about it above:

```
wait_for_completion(&kthreadd_done);
```

After this we set `gfp_allowed_mask` to `__GFP_BITS_MASK` which means that system is already running, set allowed cpus/mems to all CPUs and NUMA nodes with the `set_mems_allowed` function, allow `init` process to run on any CPU with the `set_cpus_allowed_ptr` , set pid for the `cad` or `Ctrl-Alt-Delete` , do preparation for booting of the other CPUs with the call of the `smp_prepare_cpus` , call early initcalls with the `do_pre_smp_initcalls` , initialize `SMP` with the `smp_init` and initialize lockup_detector with the call of the `lockup_detector_init` and initialize scheduler with the `sched_init_smp` .

After this we can see the call of the following functions - `do_basic_setup` . Before we will call the `do_basic_setup` function, our kernel already initialized for this moment. As comment says:

```
Now we can finally start doing some real work..
```

The `do_basic_setup` will reinitialize cpuset to the active CPUs, initialize the `khelper` - which is a kernel thread which used for making calls out to userspace from within the kernel, initialize tmpfs, initialize `drivers` subsystem, enable the user-mode helper `workqueue` and make post-early call of the `initcalls` . We can see opening of the `dev/console` and dup twice file descriptors from `0` to `2` after the `do_basic_setup` :

```
if (sys_open((const char __user *) "/dev/console", O_RDWR, 0) < 0)
    pr_err("Warning: unable to open an initial console.\n");

(void) sys_dup(0);
(void) sys_dup(0);
```

We are using two system calls here `sys_open` and `sys_dup`. In the next chapters we will see explanation and implementation of the different system calls. After we opened initial console, we check that `rdinit=` option was passed to the kernel command line or set default path of the ramdisk:

```
if (!ramdisk_execute_command)
    ramdisk_execute_command = "/init";
```

Check user's permissions for the `ramdisk` and call the `prepare_namespace` function from the init/do_mounts.c which checks and mounts the initrd:

```
if (sys_access((const char __user *) ramdisk_execute_command, 0) != 0) {
    ramdisk_execute_command = NULL;
    prepare_namespace();
}
```

This is the end of the `kernel_init_freeable` function and we need return to the `kernel_init`. The next step after the `kernel_init_freeable` finished its execution is the `async_synchronize_full`. This function waits until all asynchronous function calls have been done and after it we will call the `free_initmem` which will release all memory occupied by the initialization stuff which located between `__init_begin` and `__init_end`. After this we protect `.rodata` with the `mark_rodata_ro` and update state of the system from the `SYSTEM_BOOTING` to the

```
system_state = SYSTEM_RUNNING;
```

And tries to run the `init` process:

```
if (ramdisk_execute_command) {
    ret = run_init_process(ramdisk_execute_command);
    if (!ret)
        return 0;
    pr_err("Failed to execute %s (error %d)\n",
           ramdisk_execute_command, ret);
}
```

First of all it checks the `ramdisk_execute_command` which we set in the `kernel_init_freeable` function and it will be equal to the value of the `rdinit=` kernel command line parameters or `/init` by default. The `run_init_process` function fills the first element of the `argv_init` array:

```
static const char *argv_init[MAX_INIT_ARGS+2] = { "init", NULL, };
```

which represents arguments of the `init` program and call `do_execve` function:

```
argv_init[0] = init_filename;
return do_execve(getname_kernel(init_filename),
    (const char __user *const __user *)argv_init,
    (const char __user *const __user *)envp_init);
```

The `do_execve` function is defined in the include/linux/sched.h and runs program with the given file name and arguments. If we did not pass `rdinit=` option to the kernel command line, kernel starts to check the `execute_command` which is equal to value of the `init=` kernel command line parameter:

```
    if (execute_command) {
        ret = run_init_process(execute_command);
        if (!ret)
            return 0;
        panic("Requested init %s failed (error %d).",
```

```
                execute_command, ret);
    }
```

If we did not pass `init=` kernel command line parameter either, kernel tries to run one of the following executable files:

```
if (!try_to_run_init_process("/sbin/init") ||
    !try_to_run_init_process("/etc/init") ||
    !try_to_run_init_process("/bin/init") ||
    !try_to_run_init_process("/bin/sh"))
    return 0;
```

Otherwise we finish with panic:

```
panic("No working init found.  Try passing init= option to kernel. "
      "See Linux Documentation/init.txt for guidance.");
```

That's all! Linux kernel initialization process is finished!

Conclusion

It is the end of the tenth part about the linux kernel initialization process. It is not only the `tenth` part, but also is the last part which describes initialization of the linux kernel. As I wrote in the first part of this chapter, we will go through all steps of the kernel initialization and we did it. We started at the first architecture-independent function - `start_kernel` and finished with the launch of the first `init` process in the our system. I skipped details about different subsystem of the kernel, for example I almost did not cover scheduler, interrupts, exception handling, etc. From the next part we will start to dive to the different kernel subsystems. Hope it will be interesting.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- SLAB
- xsave
- FPU
- Documentation/security/credentials.txt
- Documentation/x86/x86_64/mm
- RCU
- VFS
- inode
- proc
- man proc
- Sysctl
- ftrace
- cgroup
- CPU hotplug documentation
- completions - wait for completion handling
- NUMA
- cpus/mems
- initcalls

* Tmpfs
* initrd
* panic
* Previous part

Interrupts and Interrupt Handling

In the following posts, we will cover interrupts and exceptions handling in the linux kernel.

- Interrupts and Interrupt Handling. Part 1. - describes interrupts and interrupt handling theory.
- Interrupts in the Linux Kernel - describes stuffs related to interrupts and exceptions handling from the early stage.
- Early interrupt handlers - describes early interrupt handlers.
- Interrupt handlers - describes first non-early interrupt handlers.
- Implementation of exception handlers - describes implementation of some exception handlers such as double fault, divide by zero etc.
- Handling non-maskable interrupts - describes handling of non-maskable interrupts and remaining interrupt handlers from the architecture-specific part.
- External hardware interrupts - describes early initialization of code which is related to handling external hardware interrupts.
- Non-early initialization of the IRQs - describes non-early initialization of code which is related to handling external hardware interrupts.
- Softirq, Tasklets and Workqueues - describes softirqs, tasklets and workqueues concepts.
- Last part - this is the last part of the `Interrupts and Interrupt Handling` chapter and here we will see a real hardware driver and some interrupts related stuff.

Interrupts and Interrupt Handling. Part 1.

Introduction

This is the first part of the new chapter of the linux insides book. We have come a long way in the previous chapter of this book. We started from the earliest steps of kernel initialization and finished with the launch of the first `init` process. Yes, we saw several initialization steps which are related to the various kernel subsystems. But we did not dig deep into the details of these subsystems. With this chapter, we will try to understand how the various kernel subsystems work and how they are implemented. As you can already understand from the chapter's title, the first subsystem will be interrupts.

What is an Interrupt?

We have already heard of the word `interrupt` in several parts of this book. We even saw a couple of examples of interrupt handlers. In the current chapter we will start from the theory i.e.,

- What are `interrupts` ?
- What are `interrupt handlers` ?

We will then continue to dig deeper into the details of `interrupts` and how the Linux kernel handles them.

The first question that arises in our mind when we come across word `interrupt` is `What is an interrupt?` An interrupt is an `event` raised by software or hardware when it needs the CPU's attention. For example, we press a button on the keyboard and what do we expect next? What should the operating system and computer do after this? To simplify matters, assume that each peripheral device has an interrupt line to the CPU. A device can use it to signal an interrupt to the CPU. However, interrupts are not signaled directly to the CPU. In the old machines there was a PIC which is a chip responsible for sequentially processing multiple interrupt requests from multiple devices. In the new machines there is an Advanced Programmable Interrupt Controller commonly known as - `APIC` . An `APIC` consists of two separate devices:

- `Local APIC`
- `I/O APIC`

The first - `Local APIC` is located on each CPU core. The local APIC is responsible for handling the CPU-specific interrupt configuration. The local APIC is usually used to manage interrupts from the APIC-timer, thermal sensor and any other such locally connected I/O devices.

The second - `I/O APIC` provides multi-processor interrupt management. It is used to distribute external interrupts among the CPU cores. More about the local and I/O APICs will be covered later in this chapter. As you can understand, interrupts can occur at any time. When an interrupt occurs, the operating system must handle it immediately. But what does it mean `to handle an interrupt` ? When an interrupt occurs, the operating system must ensure the following steps:

- The kernel must pause execution of the current process; (preempt current task);
- The kernel must search for the handler of the interrupt and transfer control (execute interrupt handler);
- After the interrupt handler completes execution, the interrupted process can resume execution.

Of course there are numerous intricacies involved in this procedure of handling interrupts. But the above 3 steps form the basic skeleton of the procedure.

Addresses of each of the interrupt handlers are maintained in a special location referred to as the - `Interrupt Descriptor Table` or `IDT` . The processor uses a unique number for recognizing the type of interruption or exception. This number is called - `vector number` . A vector number is an index in the `IDT` . There is limited amount of the vector numbers and it can be from `0` to `255` . You can note the following range-check upon the vector number within the Linux kernel source-code:

```
BUG_ON((unsigned)n > 0xFF);
```

You can find this check within the Linux kernel source code related to interrupt setup (eg. The `set_intr_gate` , `void set_system_intr_gate` in arch/x86/include/asm/desc.h). The first 32 vector numbers from `0` to `31` are reserved by the processor and used for the processing of architecture-defined exceptions and interrupts. You can find the table with the description of these vector numbers in the second part of the Linux kernel initialization process - Early interrupt and exception handling. Vector numbers from `32` to `255` are designated as user-defined interrupts and are not reserved by the processor. These interrupts are generally assigned to external I/O devices to enable those devices to send interrupts to the processor.

Now let's talk about the types of interrupts. Broadly speaking, we can split interrupts into 2 major classes:

- External or hardware generated interrupts
- Software-generated interrupts

The first - external interrupts are received through the `Local APIC` or pins on the processor which are connected to the `Local APIC` . The second - software-generated interrupts are caused by an exceptional condition in the processor itself (sometimes using special architecture-specific instructions). A common example for an exceptional condition is `division by zero` . Another example is exiting a program with the `syscall` instruction.

As mentioned earlier, an interrupt can occur at any time for a reason which the code and CPU have no control over. On the other hand, exceptions are `synchronous` with program execution and can be classified into 3 categories:

- `Faults`
- `Traps`
- `Aborts`

A `fault` is an exception reported before the execution of a "faulty" instruction (which can then be corrected). If corrected, it allows the interrupted program to be resume.

Next a `trap` is an exception which is reported immediately following the execution of the `trap` instruction. Traps also allow the interrupted program to be continued just as a `fault` does.

Finally an `abort` is an exception that does not always report the exact instruction which caused the exception and does not allow the interrupted program to be resumed.

Also we already know from the previous part that interrupts can be classified as `maskable` and `non-maskable` . Maskable interrupts are interrupts which can be blocked with the two following instructions for `x86_64` - `sti` and `cli` . We can find them in the Linux kernel source code:

```
static inline void native_irq_disable(void)
{
        asm volatile("cli": : :"memory");
}
```

and

```
static inline void native_irq_enable(void)
{
        asm volatile("sti": : :"memory");
}
```

These two instructions modify the `IF` flag bit within the interrupt register. The `sti` instruction sets the `IF` flag and the `cli` instruction clears this flag. Non-maskable interrupts are always reported. Usually any failure in the hardware is mapped to such non-maskable interrupts.

If multiple exceptions or interrupts occur at the same time, the processor handles them in order of their predefined priorities. We can determine the priorities from the highest to the lowest in the following table:

```
+----------------------------------------------------------------+
|              |                                                 |
|   Priority   | Description                                     |
|              |                                                 |
+--------------+-------------------------------------------------+
|              | Hardware Reset and Machine Checks               |
|     1        | - RESET                                         |
|              | - Machine Check                                 |
+--------------+-------------------------------------------------+
|              | Trap on Task Switch                             |
|     2        | - T flag in TSS is set                          |
|              |                                                 |
+--------------+-------------------------------------------------+
|              | External Hardware Interventions                 |
|              | - FLUSH                                         |
|     3        | - STOPCLK                                       |
|              | - SMI                                           |
|              | - INIT                                          |
+--------------+-------------------------------------------------+
|              | Traps on the Previous Instruction               |
|     4        | - Breakpoints                                   |
|              | - Debug Trap Exceptions                         |
+--------------+-------------------------------------------------+
|     5        | Nonmaskable Interrupts                          |
+--------------+-------------------------------------------------+
|     6        | Maskable Hardware Interrupts                    |
+--------------+-------------------------------------------------+
|     7        | Code Breakpoint Fault                           |
+--------------+-------------------------------------------------+
|     8        | Faults from Fetching Next Instruction           |
|              | Code-Segment Limit Violation                    |
|              | Code Page Fault                                 |
+--------------+-------------------------------------------------+
|              | Faults from Decoding the Next Instruction       |
|              | Instruction length > 15 bytes                   |
|     9        | Invalid Opcode                                  |
|              | Coprocessor Not Available                       |
|              |                                                 |
+--------------+-------------------------------------------------+
|     10       | Faults on Executing an Instruction              |
|              | Overflow                                        |
|              | Bound error                                     |
|              | Invalid TSS                                     |
|              | Segment Not Present                             |
|              | Stack fault                                     |
|              | General Protection                              |
|              | Data Page Fault                                 |
|              | Alignment Check                                 |
|              | x87 FPU Floating-point exception                |
|              | SIMD floating-point exception                   |
|              | Virtualization exception                        |
+--------------+-------------------------------------------------+
```

Now that we know a little about the various types of interrupts and exceptions, it is time to move on to a more practical part. We start with the description of the `Interrupt Descriptor Table`. As mentioned earlier, the `IDT` stores entry points of the interrupts and exceptions handlers. The `IDT` is similar in structure to the `Global Descriptor Table` which we saw in the second part of the Kernel booting process. But of course it has some differences. Instead of `descriptors`, the `IDT` entries are called `gates`. It can contain one of the following gates:

- Interrupt gates
- Task gates

* Trap gates.

in the `x86` architecture. Only long mode interrupt gates and trap gates can be referenced in the `x86_64`. Like the `Global Descriptor Table`, the `Interrupt Descriptor table` is an array of 8-byte gates on `x86` and an array of 16-byte gates on `x86_64`. We can remember from the second part of the Kernel booting process, that `Global Descriptor Table` must contain `NULL` descriptor as its first element. Unlike the `Global Descriptor Table`, the `Interrupt Descriptor Table` may contain a gate; it is not mandatory. For example, you may remember that we have loaded the Interrupt Descriptor table with the `NULL` gates only in the earlier part while transitioning into protected mode:

```
/*
 * Set up the IDT
 */
static void setup_idt(void)
{
    static const struct gdt_ptr null_idt = {0, 0};
    asm volatile("lidtl %0" : : "m" (null_idt));
}
```

from the arch/x86/boot/pm.c. The `Interrupt Descriptor table` can be located anywhere in the linear address space and the base address of it must be aligned on an 8-byte boundary on `x86` or 16-byte boundary on `x86_64`. The base address of the `IDT` is stored in the special register - `IDTR`. There are two instructions on `x86` -compatible processors to modify the `IDTR` register:

* `LIDT`
* `SIDT`

The first instruction `LIDT` is used to load the base-address of the `IDT` i.e., the specified operand into the `IDTR`. The second instruction `SIDT` is used to read and store the contents of the `IDTR` into the specified operand. The `IDTR` register is 48-bits on the `x86` and contains the following information:

```
+-----------------------------------+----------------------+
|                                   |                      |
|     Base address of the IDT       |   Limit of the IDT   |
|                                   |                      |
+-----------------------------------+----------------------+
47                                16 15                    0
```

Looking at the implementation of `setup_idt`, we have prepared a `null_idt` and loaded it to the `IDTR` register with the `lidt` instruction. Note that `null_idt` has `gdt_ptr` type which is defined as:

```
struct gdt_ptr {
        u16 len;
        u32 ptr;
} __attribute__((packed));
```

Here we can see the definition of the structure with the two fields of 2-bytes and 4-bytes each (a total of 48-bits) as we can see in the diagram. Now let's look at the `IDT` entries structure. The `IDT` entries structure is an array of the 16-byte entries which are called gates in the `x86_64`. They have the following structure:

```
127                                                                             96
+-------------------------------------------------------------------------------+
|                                                                               |
|                                Reserved                                       |
|                                                                               |
+--------------------------------------------------------------------------------
95                                                                              64
+-------------------------------------------------------------------------------+
|                                                                               |
```

```
|                              Offset 63..32                                  |
|                                                                             |
+-----------------------------------------------------------------------------+
63                               48 47      46  44   42    39             34    32
+-----------------------------------------------------------------------------+
|                                  |       |  D  |   |     |      |   |   |     |
|       Offset 31..16              |   P   |  P  | 0 |Type |0 0 0 | 0 | 0 | IST |
|                                  |       |  L  |   |     |      |   |   |     |
 -----------------------------------------------------------------------------+
31                                   16 15                                    0
+-----------------------------------------------------------------------------+
|                                      |                                      |
|          Segment Selector            |                 Offset 15..0         |
|                                      |                                      |
+-----------------------------------------------------------------------------+
```

To form an index into the IDT, the processor scales the exception or interrupt vector by sixteen. The processor handles the occurrence of exceptions and interrupts just like it handles calls of a procedure when it sees the `call` instruction. A processor uses an unique number or `vector number` of the interrupt or the exception as the index to find the necessary `Interrupt Descriptor Table` entry. Now let's take a closer look at an `IDT` entry.

As we can see, `IDT` entry on the diagram consists of the following fields:

- `0-15` bits - offset from the segment selector which is used by the processor as the base address of the entry point of the interrupt handler;
- `16-31` bits - base address of the segment select which contains the entry point of the interrupt handler;
- `IST` - a new special mechanism in the `x86_64`, will see it later;
- `DPL` - Descriptor Privilege Level;
- `P` - Segment Present flag;
- `48-63` bits - second part of the handler base address;
- `64-95` bits - third part of the base address of the handler;
- `96-127` bits - and the last bits are reserved by the CPU.

And the last `Type` field describes the type of the `IDT` entry. There are three different kinds of handlers for interrupts:

- Interrupt gate
- Trap gate
- Task gate

The `IST` or `Interrupt Stack Table` is a new mechanism in the `x86_64`. It is used as an alternative to the legacy stack-switch mechanism. Previously the `x86` architecture provided a mechanism to automatically switch stack frames in response to an interrupt. The `IST` is a modified version of the `x86` Stack switching mode. This mechanism unconditionally switches stacks when it is enabled and can be enabled for any interrupt in the `IDT` entry related with the certain interrupt (we will soon see it). From this we can understand that `IST` is not necessary for all interrupts. Some interrupts can continue to use the legacy stack switching mode. The `IST` mechanism provides up to seven `IST` pointers in the Task State Segment or `TSS` which is the special structure which contains information about a process. The `TSS` is used for stack switching during the execution of an interrupt or exception handler in the Linux kernel. Each pointer is referenced by an interrupt gate from the `IDT`.

The `Interrupt Descriptor Table` represented by the array of the `gate_desc` structures:

```
extern gate_desc idt_table[];
```

where `gate_desc` is:

```
#ifdef CONFIG_X86_64
...
...
...
```

```
typedef struct gate_struct64 gate_desc;
...
...
...
#endif
```

and `gate_struct64` defined as:

```
struct gate_struct64 {
        u16 offset_low;
        u16 segment;
        unsigned ist : 3, zero0 : 5, type : 5, dpl : 2, p : 1;
        u16 offset_middle;
        u32 offset_high;
        u32 zero1;
} __attribute__((packed));
```

Each active thread has a large stack in the Linux kernel for the `x86_64` architecture. The stack size is defined as `THREAD_SIZE` and is equal to:

```
#define PAGE_SHIFT      12
#define PAGE_SIZE       (_AC(1,UL) << PAGE_SHIFT)
...
...
...
#define THREAD_SIZE_ORDER       (2 + KASAN_STACK_ORDER)
#define THREAD_SIZE  (PAGE_SIZE << THREAD_SIZE_ORDER)
```

The `PAGE_SIZE` is `4096` -bytes and the `THREAD_SIZE_ORDER` depends on the `KASAN_STACK_ORDER` . As we can see, the `KASAN_STACK` depends on the `CONFIG_KASAN` kernel configuration parameter and is defined as:

```
#ifdef CONFIG_KASAN
    #define KASAN_STACK_ORDER 1
#else
    #define KASAN_STACK_ORDER 0
#endif
```

`KASan` is a runtime memory debugger. Thus, the `THREAD_SIZE` will be `16384` bytes if `CONFIG_KASAN` is disabled or `32768` if this kernel configuration option is enabled. These stacks contain useful data as long as a thread is alive or in a zombie state. While the thread is in user-space, the kernel stack is empty except for the `thread_info` structure (details about this structure are available in the fourth part of the Linux kernel initialization process) at the bottom of the stack. The active or zombie threads aren't the only threads with their own stack. There also exist specialized stacks that are associated with each available CPU. These stacks are active when the kernel is executing on that CPU. When the user-space is executing on the CPU, these stacks do not contain any useful information. Each CPU has a few special per-cpu stacks as well. The first is the `interrupt stack` used for the external hardware interrupts. Its size is determined as follows:

```
#define IRQ_STACK_ORDER (2 + KASAN_STACK_ORDER)
#define IRQ_STACK_SIZE (PAGE_SIZE << IRQ_STACK_ORDER)
```

or `16384` bytes. The per-cpu interrupt stack represented by the `irq_stack_union` union in the Linux kernel for `x86_64` :

```
union irq_stack_union {
    char irq_stack[IRQ_STACK_SIZE];

        struct {
        char gs_base[40];
        unsigned long stack_canary;
    };
```

```
};
```

The first `irq_stack` field is a 16 kilobytes array. Also you can see that `irq_stack_union` contains a structure with the two fields:

* `gs_base` - The `gs` register always points to the bottom of the `irqstack` union. On the `x86_64`, the `gs` register is shared by per-cpu area and stack canary (more about `per-cpu` variables you can read in the special part). All per-cpu symbols are zero based and the `gs` points to the base of the per-cpu area. You already know that segmented memory model is abolished in the long mode, but we can set the base address for the two segment registers - `fs` and `gs` with the Model specific registers and these registers can be still be used as address registers. If you remember the first part of the Linux kernel initialization process, you can remember that we have set the `gs` register:

```
    movl    $MSR_GS_BASE,%ecx
    movl    initial_gs(%rip),%eax
    movl    initial_gs+4(%rip),%edx
    wrmsr
```

where `initial_gs` points to the `irq_stack_union`:

```
GLOBAL(initial_gs)
.quad    INIT_PER_CPU_VAR(irq_stack_union)
```

* `stack_canary` - Stack canary for the interrupt stack is a `stack protector` to verify that the stack hasn't been overwritten. Note that `gs_base` is a 40 bytes array. `GCC` requires that stack canary will be on the fixed offset from the base of the `gs` and its value must be `40` for the `x86_64` and `20` for the `x86`.

The `irq_stack_union` is the first datum in the `percpu` area, we can see it in the `System.map`:

```
0000000000000000 D __per_cpu_start
0000000000000000 D irq_stack_union
0000000000004000 d exception_stacks
0000000000009000 D gdt_page
...
...
...
```

We can see its definition in the code:

```
DECLARE_PER_CPU_FIRST(union irq_stack_union, irq_stack_union) __visible;
```

Now, it's time to look at the initialization of the `irq_stack_union`. Besides the `irq_stack_union` definition, we can see the definition of the following per-cpu variables in the arch/x86/include/asm/processor.h:

```
DECLARE_PER_CPU(char *, irq_stack_ptr);
DECLARE_PER_CPU(unsigned int, irq_count);
```

The first is the `irq_stack_ptr` pointer. From the variable's name, it is obvious that this is a pointer to the top of the stack. The second - `irq_count` is used to check if a CPU is already on an interrupt stack or not. Initialization of the `irq_stack_ptr` is located in the `setup_per_cpu_areas` function in arch/x86/kernel/setup_percpu.c source code file and looks:

```
void __init setup_per_cpu_areas(void)
{
#ifdef CONFIG_X86_64
for_each_possible_cpu(cpu) {
    ...
    ...
```

```
    ...
    per_cpu(irq_stack_ptr, cpu) =
            per_cpu(irq_stack_union.irq_stack, cpu) + IRQ_STACK_SIZE;
    ...
    ...
    ...
#endif
}
```

Here we go over all the CPUs one-by-one and setup `irq_stack_ptr` .

We already know that `x86_64` has a feature called `Interrupt Stack Table` or `IST` and this feature provides the ability to switch to a new stack for events non-maskable interrupt, double fault etc. There can be up to seven `IST` entries per-cpu. Some of them are:

- `DOUBLEFAULT_STACK`
- `NMI_STACK`
- `DEBUG_STACK`
- `MCE_STACK`

or

```
#define DOUBLEFAULT_STACK 1
#define NMI_STACK 2
#define DEBUG_STACK 3
#define MCE_STACK 4
```

All interrupt-gate descriptors which switch to a new stack with the `IST` are initialized with the `set_intr_gate_ist` function. For example:

```
set_intr_gate_ist(X86_TRAP_NMI, &nmi, NMI_STACK);
...
...
...
set_intr_gate_ist(X86_TRAP_DF, &double_fault, DOUBLEFAULT_STACK);
```

where `&nmi` and `&double_fault` are addresses of the entries to the given interrupt handlers:

```
asmlinkage void nmi(void);
asmlinkage void double_fault(void);
```

defined in the [arch/x86/kernel/entry_64.S]

```
idtentry double_fault do_double_fault has_error_code=1 paranoid=2
...
...
...
ENTRY(nmi)
...
...
...
END(nmi)
```

When an interrupt or an exception occurs, the new `ss` selector is forced to `NULL` and the `ss` selector's `rpl` field is set to the new `cpl` . The old `ss` , `rsp` , register flags, `cs` , `rip` are pushed onto the new stack. In 64-bit mode, the size of interrupt stack-frame pushes is fixed at 8-bytes, so we will get the following stack:

```
+---------------+
```

```
|               |
|      SS       | 40
|      RSP      | 32
|     RFLAGS    | 24
|      CS       | 16
|      RIP      | 8
|   Error code  | 0
|               |
+---------------+
```

If the `IST` field in the interrupt gate is not `0`, we read the `IST` pointer into `rsp`. If the interrupt vector number has an error code associated with it, we then push the error code onto the stack. If the interrupt vector number has no error code, we go ahead and push the dummy error code on to the stack. We need to do this to ensure stack consistency. Next, we load the segment-selector field from the gate descriptor into the CS register and must verify that the target code-segment is a 64-bit mode code segment by the checking bit `21` i.e. the `L` bit in the `Global Descriptor Table`. Finally we load the offset field from the gate descriptor into `rip` which will be the entry-point of the interrupt handler. After this the interrupt handler begins to execute and when the interrupt handler finishes its execution, it must return control to the interrupted process with the `iret` instruction. The `iret` instruction unconditionally pops the stack pointer (`ss:rsp`) to restore the stack of the interrupted process and does not depend on the `cpl` change.

That's all.

Conclusion

It is the end of the first part of `Interrupts and Interrupt Handling` in the Linux kernel. We covered some theory and the first steps of initialization of stuffs related to interrupts and exceptions. In the next part we will continue to dive into the more practical aspects of interrupts and interrupt handling.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me a PR to linux-insides.

Links

- PIC
- Advanced Programmable Interrupt Controller
- protected mode
- long mode
- kernel stacks
- Task State Segment
- segmented memory model
- Model specific registers
- Stack canary
- Previous chapter

Interrupts and Interrupt Handling. Part 2.

Start to dive into interrupt and exceptions handling in the Linux kernel

We saw some theory about interrupts and exception handling in the previous part and as I already wrote in that part, we will start to dive into interrupts and exceptions in the Linux kernel source code in this part. As you already can note, the previous part mostly described theoretical aspects and in this part we will start to dive directly into the Linux kernel source code. We will start to do it as we did it in other chapters, from the very early places. We will not see the Linux kernel source code from the earliest code lines as we saw it for example in the Linux kernel booting process chapter, but we will start from the earliest code which is related to the interrupts and exceptions. In this part we will try to go through the all interrupts and exceptions related stuff which we can find in the Linux kernel source code.

If you've read the previous parts, you can remember that the earliest place in the Linux kernel `x86_64` architecture-specific source code which is related to the interrupt is located in the arch/x86/boot/pm.c source code file and represents the first setup of the Interrupt Descriptor Table. It occurs right before the transition into the protected mode in the `go_to_protected_mode` function by the call of the `setup_idt` :

```
void go_to_protected_mode(void)
{
    ...
    setup_idt();
    ...
}
```

The `setup_idt` function is defined in the same source code file as the `go_to_protected_mode` function and just loads the address of the `NULL` interrupts descriptor table:

```
static void setup_idt(void)
{
    static const struct gdt_ptr null_idt = {0, 0};
    asm volatile("lidtl %0" : : "m" (null_idt));
}
```

where `gdt_ptr` represents a special 48-bit `GDTR` register which must contain the base address of the `Global Descriptor Table` :

```
struct gdt_ptr {
    u16 len;
    u32 ptr;
} __attribute__((packed));
```

Of course in our case the `gdt_ptr` does not represent the `GDTR` register, but `IDTR` since we set `Interrupt Descriptor Table` . You will not find an `idt_ptr` structure, because if it had been in the Linux kernel source code, it would have been the same as `gdt_ptr` but with different name. So, as you can understand there is no sense to have two similar structures which differ only by name. You can note here, that we do not fill the `Interrupt Descriptor Table` with entries, because it is too early to handle any interrupts or exceptions at this point. That's why we just fill the `IDT` with `NULL` .

After the setup of the Interrupt descriptor table, Global Descriptor Table and other stuff we jump into protected mode in the - arch/x86/boot/pmjump.S. You can read more about it in the part which describes the transition to protected mode.

We already know from the earliest parts that entry to protected mode is located in the `boot_params.hdr.code32_start` and you can see that we pass the entry of the protected mode and `boot_params` to the `protected_mode_jump` in the end of the arch/x86/boot/pm.c:

```
protected_mode_jump(boot_params.hdr.code32_start,
                (u32)&boot_params + (ds() << 4));
```

The `protected_mode_jump` is defined in the arch/x86/boot/pmjump.S and gets these two parameters in the `ax` and `dx` registers using one of the 8086 calling conventions:

```
GLOBAL(protected_mode_jump)
    ...
    ...
    ...
    .byte    0x66, 0xea        # ljmpl opcode
2:    .long    in_pm32            # offset
    .word    __BOOT_CS        # segment
...
...
...
ENDPROC(protected_mode_jump)
```

where `in_pm32` contains a jump to the 32-bit entry point:

```
GLOBAL(in_pm32)
    ...
    ...
    jmpl    *%eax // %eax contains address of the `startup_32`
    ...
    ...
ENDPROC(in_pm32)
```

As you can remember the 32-bit entry point is in the arch/x86/boot/compressed/head_64.S assembly file, although it contains `_64` in its name. We can see the two similar files in the `arch/x86/boot/compressed` directory:

- `arch/x86/boot/compressed/head_32.S` .
- `arch/x86/boot/compressed/head_64.S` ;

But the 32-bit mode entry point is the second file in our case. The first file is not even compiled for `x86_64` . Let's look at the arch/x86/boot/compressed/Makefile:

```
vmlinux-objs-y := $(obj)/vmlinux.lds $(obj)/head_$(BITS).o $(obj)/misc.o \
...
...
```

We can see here that `head_*` depends on the `$(BITS)` variable which depends on the architecture. You can find it in the arch/x86/Makefile:

```
ifeq ($(CONFIG_X86_32),y)
...
    BITS := 32
else
    BITS := 64
    ...
endif
```

Now as we jumped on the `startup_32` from the arch/x86/boot/compressed/head_64.S we will not find anything related to the interrupt handling here. The `startup_32` contains code that makes preparations before the transition into long mode and directly jumps in to it. The `long mode` entry is located in `startup_64` and it makes preparations before the kernel decompression that occurs in the `decompress_kernel` from the arch/x86/boot/compressed/misc.c. After the kernel is decompressed, we jump on the `startup_64` from the arch/x86/kernel/head_64.S. In the `startup_64` we start to build identity-mapped pages. After we have built identity-mapped pages, checked the NX bit, setup the `Extended Feature Enable Register` (see in links), and updated the early `Global Descriptor Table` with the `lgdt` instruction, we need to setup `gs` register with the following code:

```
movl    $MSR_GS_BASE,%ecx
movl    initial_gs(%rip),%eax
movl    initial_gs+4(%rip),%edx
wrmsr
```

We already saw this code in the previous part. First of all pay attention on the last `wrmsr` instruction. This instruction writes data from the `edx:eax` registers to the model specific register specified by the `ecx` register. We can see that `ecx` contains `$MSR_GS_BASE` which is declared in the arch/x86/include/uapi/asm/msr-index.h and looks like:

```
#define MSR_GS_BASE             0xc0000101
```

From this we can understand that `MSR_GS_BASE` defines the number of the `model specific register`. Since registers `cs`, `ds`, `es`, and `ss` are not used in the 64-bit mode, their fields are ignored. But we can access memory over `fs` and `gs` registers. The model specific register provides a `back door` to the hidden parts of these segment registers and allows to use 64-bit base address for segment register addressed by the `fs` and `gs`. So the `MSR_GS_BASE` is the hidden part and this part is mapped on the `GS.base` field. Let's look on the `initial_gs`:

```
GLOBAL(initial_gs)
    .quad    INIT_PER_CPU_VAR(irq_stack_union)
```

We pass `irq_stack_union` symbol to the `INIT_PER_CPU_VAR` macro which just concatenates the `init_per_cpu__` prefix with the given symbol. In our case we will get the `init_per_cpu__irq_stack_union` symbol. Let's look at the linker script. There we can see following definition:

```
#define INIT_PER_CPU(x) init_per_cpu__##x = x + __per_cpu_load
INIT_PER_CPU(irq_stack_union);
```

It tells us that the address of the `init_per_cpu__irq_stack_union` will be `irq_stack_union + __per_cpu_load`. Now we need to understand where `init_per_cpu__irq_stack_union` and `__per_cpu_load` are what they mean. The first `irq_stack_union` is defined in the arch/x86/include/asm/processor.h with the `DECLARE_INIT_PER_CPU` macro which expands to call the `init_per_cpu_var` macro:

```
DECLARE_INIT_PER_CPU(irq_stack_union);

#define DECLARE_INIT_PER_CPU(var) \
       extern typeof(per_cpu_var(var)) init_per_cpu_var(var)

#define init_per_cpu_var(var)  init_per_cpu__##var
```

If we expand all macros we will get the same `init_per_cpu__irq_stack_union` as we got after expanding the `INIT_PER_CPU` macro, but you can note that it is not just a symbol, but a variable. Let's look at the `typeof(per_cpu_var(var))` expression. Our `var` is `irq_stack_union` and the `per_cpu_var` macro is defined in the arch/x86/include/asm/percpu.h:

```
#define PER_CPU_VAR(var)        %__percpu_seg:var
```

where:

```
#ifdef CONFIG_X86_64
    #define __percpu_seg gs
endif
```

So, we are accessing `gs:irq_stack_union` and getting its type which is `irq_union`. Ok, we defined the first variable and know its address, now let's look at the second `__per_cpu_load` symbol. There are a couple of `per-cpu` variables which are located after this symbol. The `__per_cpu_load` is defined in the include/asm-generic/sections.h:

```
extern char __per_cpu_load[], __per_cpu_start[], __per_cpu_end[];
```

and presented base address of the `per-cpu` variables from the data area. So, we know the address of the `irq_stack_union`, `__per_cpu_load` and we know that `init_per_cpu__irq_stack_union` must be placed right after `__per_cpu_load`. And we can see it in the System.map:

```
...
...
...
ffffffff819ed000 D __init_begin
ffffffff819ed000 D __per_cpu_load
ffffffff819ed000 A init_per_cpu__irq_stack_union
...
...
...
```

Now we know about `initial_gs`, so let's look at the code:

```
movl    $MSR_GS_BASE,%ecx
movl    initial_gs(%rip),%eax
movl    initial_gs+4(%rip),%edx
wrmsr
```

Here we specified a model specific register with `MSR_GS_BASE`, put the 64-bit address of the `initial_gs` to the `edx:eax` pair and execute the `wrmsr` instruction for filling the `gs` register with the base address of the `init_per_cpu__irq_stack_union` which will be at the bottom of the interrupt stack. After this we will jump to the C code on the `x86_64_start_kernel` from the arch/x86/kernel/head64.c. In the `x86_64_start_kernel` function we do the last preparations before we jump into the generic and architecture-independent kernel code and one of these preparations is filling the early `Interrupt Descriptor Table` with the interrupts handlers entries or `early_idt_handlers`. You can remember it, if you have read the part about the Early interrupt and exception handling and can remember following code:

```
for (i = 0; i < NUM_EXCEPTION_VECTORS; i++)
    set_intr_gate(i, early_idt_handlers[i]);

load_idt((const struct desc_ptr *)&idt_descr);
```

but I wrote `Early interrupt and exception handling` part when Linux kernel version was - `3.18`. For this day actual version of the Linux kernel is `4.1.0-rc6+` and `Andy Lutomirski` sent the patch and soon it will be in the mainline kernel that changes behaviour for the `early_idt_handlers`. **NOTE** While I wrote this part the patch already turned in the Linux kernel source code. Let's look on it. Now the same part looks like:

```
for (i = 0; i < NUM_EXCEPTION_VECTORS; i++)
    set_intr_gate(i, early_idt_handler_array[i]);

load_idt((const struct desc_ptr *)&idt_descr);
```

AS you can see it has only one difference in the name of the array of the interrupts handlers entry points. Now it is `early_idt_handler_arry` :

```
extern const char early_idt_handler_array[NUM_EXCEPTION_VECTORS][EARLY_IDT_HANDLER_SIZE];
```

where `NUM_EXCEPTION_VECTORS` and `EARLY_IDT_HANDLER_SIZE` are defined as:

```
#define NUM_EXCEPTION_VECTORS 32
#define EARLY_IDT_HANDLER_SIZE 9
```

So, the `early_idt_handler_array` is an array of the interrupts handlers entry points and contains one entry point on every nine bytes. You can remember that previous `early_idt_handlers` was defined in the arch/x86/kernel/head_64.S. The `early_idt_handler_array` is defined in the same source code file too:

```
ENTRY(early_idt_handler_array)
...
...
...
ENDPROC(early_idt_handler_common)
```

It fills `early_idt_handler_arry` with the `.rept NUM_EXCEPTION_VECTORS` and contains entry of the `early_make_pgtable` interrupt handler (more about its implementation you can read in the part about Early interrupt and exception handling). For now we come to the end of the `x86_64` architecture-specific code and the next part is the generic kernel code. Of course you already can know that we will return to the architecture-specific code in the `setup_arch` function and other places, but this is the end of the `x86_64` early code.

Setting stack canary for the interrupt stack

The next stop after the arch/x86/kernel/head_64.S is the biggest `start_kernel` function from the init/main.c. If you've read the previous chapter about the Linux kernel initialization process, you must remember it. This function does all initialization stuff before kernel will launch first `init` process with the pid - `1` . The first thing that is related to the interrupts and exceptions handling is the call of the `boot_init_stack_canary` function.

This function sets the canary value to protect interrupt stack overflow. We already saw a little some details about implementation of the `boot_init_stack_canary` in the previous part and now let's take a closer look on it. You can find implementation of this function in the arch/x86/include/asm/stackprotector.h and its depends on the `CONFIG_CC_STACKPROTECTOR` kernel configuration option. If this option is not set this function will not do anything:

```
#ifdef CONFIG_CC_STACKPROTECTOR
...
...
...
#else
static inline void boot_init_stack_canary(void)
{
}
#endif
```

If the `CONFIG_CC_STACKPROTECTOR` kernel configuration option is set, the `boot_init_stack_canary` function starts from the check stat `irq_stack_union` that represents per-cpu interrupt stack has offset equal to forty bytes from the `stack_canary` value:

```
#ifdef CONFIG_X86_64
        BUILD_BUG_ON(offsetof(union irq_stack_union, stack_canary) != 40);
#endif
```

As we can read in the previous part the `irq_stack_union` represented by the following union:

```
union irq_stack_union {
    char irq_stack[IRQ_STACK_SIZE];

    struct {
        char gs_base[40];
        unsigned long stack_canary;
    };
};
```

which defined in the arch/x86/include/asm/processor.h. We know that union in the C programming language is a data structure which stores only one field in a memory. We can see here that structure has first field - `gs_base` which is 40 bytes size and represents bottom of the `irq_stack`. So, after this our check with the `BUILD_BUG_ON` macro should end successfully. (you can read the first part about Linux kernel initialization process if you're interesting about the `BUILD_BUG_ON` macro).

After this we calculate new `canary` value based on the random number and Time Stamp Counter:

```
get_random_bytes(&canary, sizeof(canary));
tsc = __native_read_tsc();
canary += tsc + (tsc << 32UL);
```

and write `canary` value to the `irq_stack_union` with the `this_cpu_write` macro:

```
this_cpu_write(irq_stack_union.stack_canary, canary);
```

more about `this_cpu_*` operation you can read in the Linux kernel documentation.

Disabling/Enabling local interrupts

The next step in the init/main.c which is related to the interrupts and interrupts handling after we have set the `canary` value to the interrupt stack - is the call of the `local_irq_disable` macro.

This macro defined in the include/linux/irqflags.h header file and as you can understand, we can disable interrupts for the CPU with the call of this macro. Let's look on its implementation. First of all note that it depends on the `CONFIG_TRACE_IRQFLAGS_SUPPORT` kernel configuration option:

```
#ifdef CONFIG_TRACE_IRQFLAGS_SUPPORT
...
#define local_irq_disable() \
         do { raw_local_irq_disable(); trace_hardirqs_off(); } while (0)
...
#else
...
#define local_irq_disable()     do { raw_local_irq_disable(); } while (0)
...
#endif
```

They are both similar and as you can see have only one difference: the `local_irq_disable` macro contains call of the `trace_hardirqs_off` when `CONFIG_TRACE_IRQFLAGS_SUPPORT` is enabled. There is special feature in the lockdep subsystem - `irq-flags tracing` for tracing `hardirq` and `softirq` state. In our case `lockdep` subsystem can give us interesting information about hard/soft irqs on/off events which are occurs in the system. The `trace_hardirqs_off` function defined in the kernel/locking/lockdep.c:

```
void trace_hardirqs_off(void)
{
```

```
        trace_hardirqs_off_caller(CALLER_ADDR0);
}
EXPORT_SYMBOL(trace_hardirqs_off);
```

and just calls `trace_hardirqs_off_caller` function. The `trace_hardirqs_off_caller` checks the `hardirqs_enabled` field of the current process and increases the `redundant_hardirqs_off` if call of the `local_irq_disable` was redundant or the `hardirqs_off_events` if it was not. These two fields and other `lockdep` statistic related fields are defined in the kernel/locking/lockdep_insides.h and located in the `lockdep_stats` structure:

```
struct lockdep_stats {
...
...
...
int     softirqs_off_events;
int     redundant_softirqs_off;
...
...
...
}
```

If you will set `CONFIG_DEBUG_LOCKDEP` kernel configuration option, the `lockdep_stats_debug_show` function will write all tracing information to the `/proc/lockdep` :

```
static void lockdep_stats_debug_show(struct seq_file *m)
{
#ifdef CONFIG_DEBUG_LOCKDEP
    unsigned long long hi1 = debug_atomic_read(hardirqs_on_events),
                         hi2 = debug_atomic_read(hardirqs_off_events),
                         hr1 = debug_atomic_read(redundant_hardirqs_on),
    ...
    ...
    ...
    seq_printf(m, " hardirq on events:             %11llu\n", hi1);
    seq_printf(m, " hardirq off events:            %11llu\n", hi2);
    seq_printf(m, " redundant hardirq ons:         %11llu\n", hr1);
#endif
}
```

and you can see its result with the:

```
$ sudo cat /proc/lockdep
 hardirq on events:             12838248974
 hardirq off events:            12838248979
 redundant hardirq ons:               67792
 redundant hardirq offs:         3836339146
 softirq on events:                38002159
 softirq off events:               38002187
 redundant softirq ons:                   0
 redundant softirq offs:                  0
```

Ok, now we know a little about tracing, but more info will be in the separate part about `lockdep` and `tracing` . You can see that the both `local_disable_irq` macros have the same part - `raw_local_irq_disable` . This macro defined in the arch/x86/include/asm/irqflags.h and expands to the call of the:

```
static inline void native_irq_disable(void)
{
        asm volatile("cli": : :"memory");
}
```

And you already must remember that `cli` instruction clears the IF flag which determines ability of a processor to handle an interrupt or an exception. Besides the `local_irq_disable`, as you already can know there is an inverse macro - `local_irq_enable`. This macro has the same tracing mechanism and very similar on the `local_irq_enable`, but as you can understand from its name, it enables interrupts with the `sti` instruction:

```
static inline void native_irq_enable(void)
{
        asm volatile("sti": : :"memory");
}
```

Now we know how `local_irq_disable` and `local_irq_enable` work. It was the first call of the `local_irq_disable` macro, but we will meet these macros many times in the Linux kernel source code. But for now we are in the `start_kernel` function from the init/main.c and we just disabled `local` interrupts. Why local and why we did it? Previously kernel provided a method to disable interrupts on all processors and it was called `cli`. This function was removed and now we have `local_irq_{enabled,disable}` to disable or enable interrupts on the current processor. After we've disabled the interrupts with the `local_irq_disable` macro, we set the:

```
early_boot_irqs_disabled = true;
```

The `early_boot_irqs_disabled` variable defined in the include/linux/kernel.h:

```
extern bool early_boot_irqs_disabled;
```

and used in the different places. For example it used in the `smp_call_function_many` function from the kernel/smp.c for the checking possible deadlock when interrupts are disabled:

```
WARN_ON_ONCE(cpu_online(this_cpu) && irqs_disabled()
                     && !oops_in_progress && !early_boot_irqs_disabled);
```

Early trap initialization during kernel initialization

The next functions after the `local_disable_irq` are `boot_cpu_init` and `page_address_init`, but they are not related to the interrupts and exceptions (more about this functions you can read in the chapter about Linux kernel initialization process). The next is the `setup_arch` function. As you can remember this function located in the arch/x86/kernel/setup.c source code file and makes initialization of many different architecture-dependent stuff. The first interrupts related function which we can see in the `setup_arch` is the - `early_trap_init` function. This function defined in the arch/x86/kernel/traps.c and fills `Interrupt Descriptor Table` with the couple of entries:

```
void __init early_trap_init(void)
{
        set_intr_gate_ist(X86_TRAP_DB, &debug, DEBUG_STACK);
        set_system_intr_gate_ist(X86_TRAP_BP, &int3, DEBUG_STACK);
#ifdef CONFIG_X86_32
        set_intr_gate(X86_TRAP_PF, page_fault);
#endif
        load_idt(&idt_descr);
}
```

Here we can see calls of three different functions:

- `set_intr_gate_ist`
- `set_system_intr_gate_ist`
- `set_intr_gate`

All of these functions defined in the arch/x86/include/asm/desc.h and do the similar thing but not the same. The first `set_intr_gate_ist` function inserts new an interrupt gate in the `IDT`. Let's look on its implementation:

```
static inline void set_intr_gate_ist(int n, void *addr, unsigned ist)
{
        BUG_ON((unsigned)n > 0xFF);
        _set_gate(n, GATE_INTERRUPT, addr, 0, ist, __KERNEL_CS);
}
```

First of all we can see the check that `n` which is vector number of the interrupt is not greater than `0xff` or 255. We need to check it because we remember from the previous part that vector number of an interrupt must be between `0` and `255`. In the next step we can see the call of the `_set_gate` function that sets a given interrupt gate to the `IDT` table:

```
static inline void _set_gate(int gate, unsigned type, void *addr,
                             unsigned dpl, unsigned ist, unsigned seg)
{
        gate_desc s;

        pack_gate(&s, type, (unsigned long)addr, dpl, ist, seg);
        write_idt_entry(idt_table, gate, &s);
        write_trace_idt_entry(gate, &s);
}
```

Here we start from the `pack_gate` function which takes clean `IDT` entry represented by the `gate_desc` structure and fills it with the base address and limit, Interrupt Stack Table, Privilege level, type of an interrupt which can be one of the following values:

- `GATE_INTERRUPT`
- `GATE_TRAP`
- `GATE_CALL`
- `GATE_TASK`

and set the present bit for the given `IDT` entry:

```
static inline void pack_gate(gate_desc *gate, unsigned type, unsigned long func,
                             unsigned dpl, unsigned ist, unsigned seg)
{
        gate->offset_low        = PTR_LOW(func);
        gate->segment           = __KERNEL_CS;
        gate->ist               = ist;
        gate->p                 = 1;
        gate->dpl               = dpl;
        gate->zero0             = 0;
        gate->zero1             = 0;
        gate->type              = type;
        gate->offset_middle     = PTR_MIDDLE(func);
        gate->offset_high       = PTR_HIGH(func);
}
```

After this we write just filled interrupt gate to the `IDT` with the `write_idt_entry` macro which expands to the `native_write_idt_entry` and just copy the interrupt gate to the `idt_table` table by the given index:

```
#define write_idt_entry(dt, entry, g)           native_write_idt_entry(dt, entry, g)

static inline void native_write_idt_entry(gate_desc *idt, int entry, const gate_desc *gate)
{
        memcpy(&idt[entry], gate, sizeof(*gate));
}
```

where `idt_table` is just array of `gate_desc` :

```
extern gate_desc idt_table[];
```

That's all. The second `set_system_intr_gate_ist` function has only one difference from the `set_intr_gate_ist` :

```
static inline void set_system_intr_gate_ist(int n, void *addr, unsigned ist)
{
        BUG_ON((unsigned)n > 0xFF);
        _set_gate(n, GATE_INTERRUPT, addr, 0x3, ist, __KERNEL_CS);
}
```

Do you see it? Look on the fourth parameter of the `_set_gate` . It is `0x3` . In the `set_intr_gate` it was `0x0` . We know that this parameter represent `DPL` or privilege level. We also know that `0` is the highest privilege level and `3` is the lowest.Now we know how `set_system_intr_gate_ist` , `set_intr_gate_ist` , `set_intr_gate` are work and we can return to the `early_trap_init` function. Let's look on it again:

```
set_intr_gate_ist(X86_TRAP_DB, &debug, DEBUG_STACK);
set_system_intr_gate_ist(X86_TRAP_BP, &int3, DEBUG_STACK);
```

We set two `IDT` entries for the `#DB` interrupt and `int3` . These functions takes the same set of parameters:

- vector number of an interrupt;
- address of an interrupt handler;
- interrupt stack table index.

That's all. More about interrupts and handlers you will know in the next parts.

Conclusion

It is the end of the second part about interrupts and interrupt handling in the Linux kernel. We saw the some theory in the previous part and started to dive into interrupts and exceptions handling in the current part. We have started from the earliest parts in the Linux kernel source code which are related to the interrupts. In the next part we will continue to dive into this interesting theme and will know more about interrupt handling process.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- IDT
- Protected mode
- List of x86 calling conventions
- 8086
- Long mode
- NX
- Extended Feature Enable Register
- Model-specific register
- Process identifier
- lockdep

- irqflags tracing
- IF
- Stack canary
- Union type
- this*cpu** operations
- vector number
- Interrupt Stack Table
- Privilege level
- Previous part

Interrupts and Interrupt Handling. Part 3.

Exception Handling

This is the third part of the chapter about an interrupts and an exceptions handling in the Linux kernel and in the previous part we stopped at the `setup_arch` function from the arch/x86/kernel/setup.c source code file.

We already know that this function executes initialization of architecture-specific stuff. In our case the `setup_arch` function does x86_64 architecture related initializations. The `setup_arch` is big function, and in the previous part we stopped on the setting of the two exceptions handlers for the two following exceptions:

- `#DB` - debug exception, transfers control from the interrupted process to the debug handler;
- `#BP` - breakpoint exception, caused by the `int 3` instruction.

These exceptions allow the `x86_64` architecture to have early exception processing for the purpose of debugging via the kgdb.

As you can remember we set these exceptions handlers in the `early_trap_init` function:

```
void __init early_trap_init(void)
{
        set_intr_gate_ist(X86_TRAP_DB, &debug, DEBUG_STACK);
        set_system_intr_gate_ist(X86_TRAP_BP, &int3, DEBUG_STACK);
        load_idt(&idt_descr);
}
```

from the arch/x86/kernel/traps.c. We already saw implementation of the `set_intr_gate_ist` and `set_system_intr_gate_ist` functions in the previous part and now we will look on the implementation of these two exceptions handlers.

Debug and Breakpoint exceptions

Ok, we setup exception handlers in the `early_trap_init` function for the `#DB` and `#BP` exceptions and now time is to consider their implementations. But before we will do this, first of all let's look on details of these exceptions.

The first exceptions - `#DB` or `debug` exception occurs when a debug event occurs. For example - attempt to change the contents of a debug register. Debug registers are special registers that were presented in `x86` processors starting from the Intel 80386 processor and as you can understand from name of this CPU extension, main purpose of these registers is debugging.

These registers allow to set breakpoints on the code and read or write data to trace it. Debug registers may be accessed only in the privileged mode and an attempt to read or write the debug registers when executing at any other privilege level causes a general protection fault exception. That's why we have used `set_intr_gate_ist` for the `#DB` exception, but not the `set_system_intr_gate_ist`.

The verctor number of the `#DB` exceptions is `1` (we pass it as `X86_TRAP_DB`) and as we may read in specification, this exception has no error code:

```
+-----------------------------------------------------+
|Vector|Mnemonic|Description         |Type |Error Code|
+-----------------------------------------------------+
|1     | #DB    |Reserved            |F/T  |NO        |
+-----------------------------------------------------+
```

The second exception is `#BP` or `breakpoint` exception occurs when processor executes the int 3 instruction. Unlike the `DB` exception, the `#BP` exception may occur in userspace. We can add it anywhere in our code, for example let's look on the simple program:

```
// breakpoint.c
#include <stdio.h>

int main() {
    int i;
    while (i < 6){
        printf("i equal to: %d\n", i);
        __asm__("int3");
        ++i;
    }
}
```

If we will compile and run this program, we will see following output:

```
$ gcc breakpoint.c -o breakpoint
i equal to: 0
Trace/breakpoint trap
```

But if will run it with gdb, we will see our breakpoint and can continue execution of our program:

```
$ gdb breakpoint
...
...
...
(gdb) run
Starting program: /home/alex/breakpoints
i equal to: 0

Program received signal SIGTRAP, Trace/breakpoint trap.
0x0000000000400585 in main ()
=> 0x0000000000400585 <main+31>:    83 45 fc 01    add    DWORD PTR [rbp-0x4],0x1
(gdb) c
Continuing.
i equal to: 1

Program received signal SIGTRAP, Trace/breakpoint trap.
0x0000000000400585 in main ()
=> 0x0000000000400585 <main+31>:    83 45 fc 01    add    DWORD PTR [rbp-0x4],0x1
(gdb) c
Continuing.
i equal to: 2

Program received signal SIGTRAP, Trace/breakpoint trap.
0x0000000000400585 in main ()
=> 0x0000000000400585 <main+31>:    83 45 fc 01    add    DWORD PTR [rbp-0x4],0x1
...
...
...
```

From this moment we know a little about these two exceptions and we can move on to consideration of their handlers.

Preparation before an exception handler

As you may note before, the `set_intr_gate_ist` and `set_system_intr_gate_ist` functions takes an addresses of exceptions handlers in theirs second parameter. In or case our two exception handlers will be:

* `debug` ;
* `int3` .

You will not find these functions in the C code. all of that could be found in the kernel's `*.c/*.h` files only definition of these functions which are located in the arch/x86/include/asm/traps.h kernel header file:

```
asmlinkage void debug(void);
```

and

```
asmlinkage void int3(void);
```

You may note `asmlinkage` directive in definitions of these functions. The directive is the special specificator of the gcc. Actually for a `C` functions which are called from assembly, we need in explicit declaration of the function calling convention. In our case, if function made with `asmlinkage` descriptor, then `gcc` will compile the function to retrieve parameters from stack.

So, both handlers are defined in the arch/x86/entry/entry_64.S assembly source code file with the `idtentry` macro:

```
idtentry debug do_debug has_error_code=0 paranoid=1 shift_ist=DEBUG_STACK
```

and

```
idtentry int3 do_int3 has_error_code=0 paranoid=1 shift_ist=DEBUG_STACK
```

Each exception handler may be consists from two parts. The first part is generic part and it is the same for all exception handlers. An exception handler should to save general purpose registers on the stack, switch to kernel stack if an exception came from userspace and transfer control to the second part of an exception handler. The second part of an exception handler does certain work depends on certain exception. For example page fault exception handler should find virtual page for given address, invalid opcode exception handler should send `SIGILL` signal and etc.

As we just saw, an exception handler starts from definition of the `idtentry` macro from the arch/x86/kernel/entry_64.S assembly source code file, so let's look at implementation of this macro. As we may see, the `idtentry` macro takes five arguments:

* `sym` - defines global symbol with the `.globl name` which will be an an entry of exception handler;
* `do_sym` - symbol name which represents a secondary entry of an exception handler;
* `has_error_code` - information about existence of an error code of exception.

The last two parameters are optional:

* `paranoid` - shows us how we need to check current mode (will see explanation in details later);
* `shift_ist` - shows us is an exception running at `Interrupt Stack Table` .

Definition of the `.idtentry` macro looks:

```
.macro idtentry sym do_sym has_error_code:req paranoid=0 shift_ist=-1
ENTRY(\sym)
...
...
...
END(\sym)
.endm
```

Before we will consider internals of the `idtentry` macro, we should to know state of stack when an exception occurs. As we may read in the Intel® 64 and IA-32 Architectures Software Developer's Manual 3A, the state of stack when an exception occurs is following:

```
    +------------+
+40 | %SS        |
+32 | %RSP       |
+24 | %RFLAGS    |
+16 | %CS        |
 +8 | %RIP       |
  0 | ERROR CODE | <-- %RSP
    +------------+
```

Now we may start to consider implementation of the `idtmacro`. Both `#DB` and `BP` exception handlers are defined as:

```
idtentry debug do_debug has_error_code=0 paranoid=1 shift_ist=DEBUG_STACK
idtentry int3 do_int3 has_error_code=0 paranoid=1 shift_ist=DEBUG_STACK
```

If we will look at these definitions, we may know that compiler will generate two routines with `debug` and `int3` names and both of these exception handlers will call `do_debug` and `do_int3` secondary handlers after some preparation. The third parameter defines existence of error code and as we may see both our exception do not have them. As we may see on the diagram above, processor pushes error code on stack if an exception provides it. In our case, the `debug` and `int3` exception do not have error codes. This may bring some difficulties because stack will look differently for exceptions which provides error code and for exceptions which not. That's why implementation of the `idtentry` macro starts from putting a fake error code to the stack if an exception does not provide it:

```
.ifeq \has_error_code
    pushq    $-1
.endif
```

But it is not only fake error-code. Moreover the `-1` also represents invalid system call number, so that the system call restart logic will not be triggered.

The last two parameters of the `idtentry` macro `shift_ist` and `paranoid` allow to know do an exception handler runned at stack from `Interrupt Stack Table` or not. You already may know that each kernel thread in the system has own stack. In addition to these stacks, there are some specialized stacks associated with each processor in the system. One of these stacks is - exception stack. The x86_64 architecture provides special feature which is called - `Interrupt Stack Table`. This feature allows to switch to a new stack for designated events such as an atomic exceptions like `double fault` and etc. So the `shift_ist` parameter allows us to know do we need to switch on `IST` stack for an exception handler or not.

The second parameter - `paranoid` defines the method which helps us to know did we come from userspace or not to an exception handler. The easiest way to determine this is to via `CPL` or `Current Privilege Level` in `CS` segment register. If it is equal to `3`, we came from userspace, if zero we came from kernel space:

```
testl $3,CS(%rsp)
jnz userspace
...
...
...
// we are from the kernel space
```

But unfortunately this method does not give a 100% guarantee. As described in the kernel documentation:

> if we are in an NMI/MCE/DEBUG/whatever super-atomic entry context, which might have triggered right after a normal entry wrote CS to the stack but before we executed SWAPGS, then the only safe way to check for GS is the slower method: the RDMSR.

In other words for example `NMI` could happen inside the critical section of a swapgs instruction. In this way we should check value of the `MSR_GS_BASE` model specific register which stores pointer to the start of per-cpu area. So to check did we come from userspace or not, we should to check value of the `MSR_GS_BASE` model specific register and if it is negative we came from kernel space, in other way we came from userspace:

```
movl $MSR_GS_BASE,%ecx
rdmsr
testl %edx,%edx
js 1f
```

In first two lines of code we read value of the `MSR_GS_BASE` model specific register into `edx:eax` pair. We can't set negative value to the `gs` from userspace. But from other side we know that direct mapping of the physical memory starts from the `0xffff880000000000` virtual address. In this way, `MSR_GS_BASE` will contain an address from `0xffff880000000000` to `0xffffc7ffffffffff`. After the `rdmsr` instruction will be executed, the smallest possible value in the `%edx` register will be - `0xffff8800` which is `-30720` in unsigned 4 bytes. That's why kernel space `gs` which points to start of `per-cpu` area will contain negative value.

After we pushed fake error code on the stack, we should allocate space for general purpose registers with:

```
ALLOC_PT_GPREGS_ON_STACK
```

macro which is defined in the arch/x86/entry/calling.h header file. This macro just allocates 15*8 bytes space on the stack to preserve general purpose registers:

```
.macro ALLOC_PT_GPREGS_ON_STACK addskip=0
    addq    $-(15*8+\addskip), %rsp
.endm
```

So the stack will look like this after execution of the `ALLOC_PT_GPREGS_ON_STACK`:

```
     +------------+
+160 | %SS        |
+152 | %RSP       |
+144 | %RFLAGS    |
+136 | %CS        |
+128 | %RIP       |
+120 | ERROR CODE |
     |------------|
+112 |            |
+104 |            |
 +96 |            |
 +88 |            |
 +80 |            |
 +72 |            |
 +64 |            |
 +56 |            |
 +48 |            |
 +40 |            |
 +32 |            |
 +24 |            |
 +16 |            |
  +8 |            |
  +0 |            | <- %RSP
     +------------+
```

After we allocated space for general purpose registers, we do some checks to understand did an exception come from userspace or not and if yes, we should move back to an interrupted process stack or stay on exception stack:

```
.if \paranoid
    .if \paranoid == 1
        testb   $3, CS(%rsp)
        jnz     1f
    .endif
    call    paranoid_entry
.else
    call    error_entry
.endif
```

Let's consider all of these there cases in course.

An exception occured in userspace

In the first let's consider a case when an exception has `paranoid=1` like our `debug` and `int3` exceptions. In this case we check selector from `CS` segment register and jump at `1f` label if we came from userspace or the `paranoid_entry` will be called in other way.

Let's consider first case when we came from userspace to an exception handler. As described above we should jump at `1` label. The `1` label starts from the call of the

```
call    error_entry
```

routine which saves all general purpose registers in the previously allocated area on the stack:

```
SAVE_C_REGS 8
SAVE_EXTRA_REGS 8
```

These both macros are defined in the arch/x86/entry/calling.h header file and just move values of general purpose registers to a certain place at the stack, for example:

```
.macro SAVE_EXTRA_REGS offset=0
    movq %r15, 0*8+\offset(%rsp)
    movq %r14, 1*8+\offset(%rsp)
    movq %r13, 2*8+\offset(%rsp)
    movq %r12, 3*8+\offset(%rsp)
    movq %rbp, 4*8+\offset(%rsp)
    movq %rbx, 5*8+\offset(%rsp)
.endm
```

After execution of `SAVE_C_REGS` and `SAVE_EXTRA_REGS` the stack will look:

```
     +------------+
+160 | %SS        |
+152 | %RSP       |
+144 | %RFLAGS    |
+136 | %CS        |
+128 | %RIP       |
+120 | ERROR CODE |
     |------------|
+112 | %RDI       |
+104 | %RSI       |
 +96 | %RDX       |
 +88 | %RCX       |
 +80 | %RAX       |
 +72 | %R8        |
 +64 | %R9        |
 +56 | %R10       |
```

```
  +48 | %R11        |
  +40 | %RBX        |
  +32 | %RBP        |
  +24 | %R12        |
  +16 | %R13        |
   +8 | %R14        |
   +0 | %R15        | <- %RSP
      +-------------+
```

After the kernel saved general purpose registers at the stack, we should check that we came from userspace space again with:

```
testb    $3, CS+8(%rsp)
jz    .Lerror_kernelspace
```

because we may have potentially fault if as described in documentation truncated `%RIP` was reported. Anyway, in both cases the SWAPGS instruction will be executed and values from `MSR_KERNEL_GS_BASE` and `MSR_GS_BASE` will be swapped. From this moment the `%gs` register will point to the base address of kernel structures. So, the `SWAPGS` instruction is called and it was main point of the `error_entry` routing.

Now we can back to the `idtentry` macro. We may see following assembler code after the call of `error_entry` :

```
movq    %rsp, %rdi
call    sync_regs
```

Here we put base address of stack pointer `%rdi` register which will be first argument (according to x86_64 ABI) of the `sync_regs` function and call this function which is defined in the arch/x86/kernel/traps.c source code file:

```
asmlinkage __visible notrace struct pt_regs *sync_regs(struct pt_regs *eregs)
{
    struct pt_regs *regs = task_pt_regs(current);
    *regs = *eregs;
    return regs;
}
```

This function takes the result of the `task_ptr_regs` macro which is defined in the arch/x86/include/asm/processor.h header file, stores it in the stack pointer and return it. The `task_ptr_regs` macro expands to the address of `thread.sp0` which represents pointer to the normal kernel stack:

```
#define task_pt_regs(tsk)       ((struct pt_regs *)(tsk)->thread.sp0 - 1)
```

As we came from userspace, this means that exception handler will run in real process context. After we got stack pointer from the `sync_regs` we switch stack:

```
movq    %rax, %rsp
```

The last two steps before an exception handler will call secondary handler are:

1. Passing pointer to `pt_regs` structure which contains preserved general purpose registers to the `%rdi` register:

```
movq    %rsp, %rdi
```

as it will be passed as first parameter of secondary exception handler.

1. Pass error code to the `%rsi` register as it will be second argument of an exception handler and set it to `-1` on the stack for the same purpose as we did it before - to prevent restart of a system call:

```
.if \has_error_code
    movq    ORIG_RAX(%rsp), %rsi
    movq    $-1, ORIG_RAX(%rsp)
.else
    xorl    %esi, %esi
.endif
```

Additionally you may see that we zeroed the `%esi` register above in a case if an exception does not provide error code.

In the end we just call secondary exception handler:

```
call    \do_sym
```

which:

```
dotraplinkage void do_debug(struct pt_regs *regs, long error_code);
```

will be for `debug` exception and:

```
dotraplinkage void notrace do_int3(struct pt_regs *regs, long error_code);
```

will be for `int 3` exception. In this part we will not see implementations of secondary handlers, because of they are very specific, but will see some of them in one of next parts.

We just considered first case when an exception occurred in userspace. Let's consider last two.

An exception with paranoid > 0 occurred in kernelspace

In this case an exception was occurred in kernelspace and `idtentry` macro is defined with `paranoid=1` for this exception. This value of `paranoid` means that we should use slower way that we saw in the beginning of this part to check do we really came from kernelspace or not. The `paranoid_entry` routing allows us to know this:

```
ENTRY(paranoid_entry)
    cld
    SAVE_C_REGS 8
    SAVE_EXTRA_REGS 8
    movl    $1, %ebx
    movl    $MSR_GS_BASE, %ecx
    rdmsr
    testl    %edx, %edx
    js    1f
    SWAPGS
    xorl    %ebx, %ebx
1:    ret
END(paranoid_entry)
```

As you may see, this function represents the same that we covered before. We use second (slow) method to get information about previous state of an interrupted task. As we checked this and executed `SWAPGS` in a case if we came from userspace, we should to do the same that we did before: We need to put pointer to a structure which holds general purpose registers to the `%rdi` (which will be first parameter of a secondary handler) and put error code if an exception provides it to the `%rsi` (which will be second parameter of a secondary handler):

```
movq    %rsp, %rdi

.if \has_error_code
    movq    ORIG_RAX(%rsp), %rsi
```

```
    movq    $-1, ORIG_RAX(%rsp)
.else
    xorl    %esi, %esi
.endif
```

The last step before a secondary handler of an exception will be called is cleanup of new `IST` stack fram:

```
.if \shift_ist != -1
    subq    $EXCEPTION_STKSZ, CPU_TSS_IST(\shift_ist)
.endif
```

You may remember that we passed the `shift_ist` as argument of the `idtentry` macro. Here we check its value and if its not equal to `-1`, we get pointer to a stack from `Interrupt Stack Table` by `shift_ist` index and setup it.

In the end of this second way we just call secondary exception handler as we did it before:

```
call    \do_sym
```

The last method is similar to previous both, but an exception occured with `paranoid=0` and we may use fast method determination of where we are from.

Exit from an exception handler

After secondary handler will finish its works, we will return to the `idtentry` macro and the next step will be jump to the `error_exit`:

```
jmp    error_exit
```

routine. The `error_exit` function defined in the same arch/x86/entry/entry_64.S assembly source code file and the main goal of this function is to know where we are from (from userspace or kernelspace) and execute `SWPAGS` depends on this. Restore registers to previous state and execute `iret` instruction to transfer control to an interrupted task.

That's all.

Conclusion

It is the end of the third part about interrupts and interrupt handling in the Linux kernel. We saw the initialization of the Interrupt descriptor table in the previous part with the `#DB` and `#BP` gates and started to dive into preparation before control will be transferred to an exception handler and implementation of some interrupt handlers in this part. In the next part we will continue to dive into this theme and will go next by the `setup_arch` function and will try to understand interrupts handling related stuff.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- Debug registers
- Intel 80385
- INT 3
- gcc

- TSS
- GNU assembly .error directive
- dwarf2
- CFI directives
- IRQ
- system call
- swapgs
- SIGTRAP
- Per-CPU variables
- kgdb
- ACPI
- Previous part

Interrupts and Interrupt Handling. Part 4.

Initialization of non-early interrupt gates

This is fourth part about an interrupts and exceptions handling in the Linux kernel and in the previous part we saw first early `#DB` and `#BP` exceptions handlers from the arch/x86/kernel/traps.c. We stopped on the right after the `early_trap_init` function that called in the `setup_arch` function which defined in the arch/x86/kernel/setup.c. In this part we will continue to dive into an interrupts and exceptions handling in the Linux kernel for `x86_64` and continue to do it from the place where we left off in the last part. First thing which is related to the interrupts and exceptions handling is the setup of the `#PF` or page fault handler with the `early_trap_pf_init` function. Let's start from it.

Early page fault handler

The `early_trap_pf_init` function defined in the arch/x86/kernel/traps.c. It uses `set_intr_gate` macro that fills Interrupt Descriptor Table with the given entry:

```
void __init early_trap_pf_init(void)
{
#ifdef CONFIG_X86_64
         set_intr_gate(X86_TRAP_PF, page_fault);
#endif
}
```

This macro defined in the arch/x86/include/asm/desc.h. We already saw macros like this in the previous part - `set_system_intr_gate` and `set_intr_gate_ist`. This macro checks that given vector number is not greater than `255` (maximum vector number) and calls `_set_gate` function as `set_system_intr_gate` and `set_intr_gate_ist` did it:

```
#define set_intr_gate(n, addr)                                  \
do {                                                            \
        BUG_ON((unsigned)n > 0xFF);                             \
        _set_gate(n, GATE_INTERRUPT, (void *)addr, 0, 0,        \
                  __KERNEL_CS);                                 \
        _trace_set_gate(n, GATE_INTERRUPT, (void *)trace_##addr,\
                        0, 0, __KERNEL_CS);                     \
} while (0)
```

The `set_intr_gate` macro takes two parameters:

- vector number of a interrupt;
- address of an interrupt handler;

In our case they are:

- `X86_TRAP_PF` - `14`;
- `page_fault` - the interrupt handler entry point.

The `X86_TRAP_PF` is the element of enum which defined in the arch/x86/include/asm/traprs.h:

```
enum {
    ...
    ...
    ...
    ...
    X86_TRAP_PF,            /* 14, Page Fault */
```

```
    ...
    ...
    ...
}
```

When the `early_trap_pf_init` will be called, the `set_intr_gate` will be expanded to the call of the `_set_gate` which will fill the `IDT` with the handler for the page fault. Now let's look on the implementation of the `page_fault` handler. The `page_fault` handler defined in the arch/x86/kernel/entry_64.S assembly source code file as all exceptions handlers. Let's look on it:

```
trace_idtentry page_fault do_page_fault has_error_code=1
```

We saw in the previous part how `#DB` and `#BP` handlers defined. They were defined with the `idtentry` macro, but here we can see `trace_idtentry`. This macro defined in the same source code file and depends on the `CONFIG_TRACING` kernel configuration option:

```
#ifdef CONFIG_TRACING
.macro trace_idtentry sym do_sym has_error_code:req
idtentry trace(\sym) trace(\do_sym) has_error_code=\has_error_code
idtentry \sym \do_sym has_error_code=\has_error_code
.endm
#else
.macro trace_idtentry sym do_sym has_error_code:req
idtentry \sym \do_sym has_error_code=\has_error_code
.endm
#endif
```

We will not dive into exceptions Tracing now. If `CONFIG_TRACING` is not set, we can see that `trace_idtentry` macro just expands to the normal `idtentry`. We already saw implementation of the `idtentry` macro in the previous part, so let's start from the `page_fault` exception handler.

As we can see in the `idtentry` definition, the handler of the `page_fault` is `do_page_fault` function which defined in the arch/x86/mm/fault.c and as all exceptions handlers it takes two arguments:

- `regs` - `pt_regs` structure that holds state of an interrupted process;
- `error_code` - error code of the page fault exception.

Let's look inside this function. First of all we read content of the cr2 control register:

```
dotraplinkage void notrace
do_page_fault(struct pt_regs *regs, unsigned long error_code)
{
    unsigned long address = read_cr2();
    ...
    ...
    ...
}
```

This register contains a linear address which caused `page fault`. In the next step we make a call of the `exception_enter` function from the include/linux/context_tracking.h. The `exception_enter` and `exception_exit` are functions from context tracking subsystem in the Linux kernel used by the RCU to remove its dependency on the timer tick while a processor runs in userspace. Almost in the every exception handler we will see similar code:

```
enum ctx_state prev_state;
prev_state = exception_enter();
...
... // exception handler here
...
```

```
exception_exit(prev_state);
```

The `exception_enter` function checks that `context tracking` is enabled with the `context_tracking_is_enabled` and if it is in enabled state, we get previous context with the `this_cpu_read` (more about `this_cpu_*` operations you can read in the Documentation). After this it calls `context_tracking_user_exit` function which informs the context tracking that the processor is exiting userspace mode and entering the kernel:

```
static inline enum ctx_state exception_enter(void)
{
        enum ctx_state prev_ctx;

        if (!context_tracking_is_enabled())
                return 0;

        prev_ctx = this_cpu_read(context_tracking.state);
        context_tracking_user_exit();

        return prev_ctx;
}
```

The state can be one of the:

```
enum ctx_state {
    IN_KERNEL = 0,
    IN_USER,
} state;
```

And in the end we return previous context. Between the `exception_enter` and `exception_exit` we call actual page fault handler:

```
__do_page_fault(regs, error_code, address);
```

The `__do_page_fault` is defined in the same source code file as `do_page_fault` - arch/x86/mm/fault.c. In the beginning of the `__do_page_fault` we check state of the kmemcheck checker. The `kmemcheck` detects warns about some uses of uninitialized memory. We need to check it because page fault can be caused by kmemcheck:

```
if (kmemcheck_active(regs))
        kmemcheck_hide(regs);
    prefetchw(&mm->mmap_sem);
```

After this we can see the call of the `prefetchw` which executes instruction with the same name which fetches X86_FEATURE_3DNOW to get exclusive cache line. The main purpose of prefetching is to hide the latency of a memory access. In the next step we check that we got page fault not in the kernel space with the following condition:

```
if (unlikely(fault_in_kernel_space(address))) {
...
...
...
}
```

where `fault_in_kernel_space` is:

```
static int fault_in_kernel_space(unsigned long address)
{
        return address >= TASK_SIZE_MAX;
}
```

The `TASK_SIZE_MAX` macro expands to the:

```
#define TASK_SIZE_MAX   ((1UL << 47) - PAGE_SIZE)
```

or `0x00007ffffffff000` . Pay attention on `unlikely` macro. There are two macros in the Linux kernel:

```
#define likely(x)      __builtin_expect(!!(x), 1)
#define unlikely(x)    __builtin_expect(!!(x), 0)
```

You can often find these macros in the code of the Linux kernel. Main purpose of these macros is optimization. Sometimes this situation is that we need to check the condition of the code and we know that it will rarely be `true` or `false` . With these macros we can tell to the compiler about this. For example

```
static int proc_root_readdir(struct file *file, struct dir_context *ctx)
{
        if (ctx->pos < FIRST_PROCESS_ENTRY) {
                int error = proc_readdir(file, ctx);
                if (unlikely(error <= 0))
                        return error;
...
...
...
}
```

Here we can see `proc_root_readdir` function which will be called when the Linux VFS needs to read the `root` directory contents. If condition marked with `unlikely` , compiler can put `false` code right after branching. Now let's back to the our address check. Comparison between the given address and the `0x00007ffffffff000` will give us to know, was page fault in the kernel mode or user mode. After this check we know it. After this `__do_page_fault` routine will try to understand the problem that provoked page fault exception and then will pass address to the appropriate routine. It can be `kmemcheck` fault, spurious fault, kprobes fault and etc. Will not dive into implementation details of the page fault exception handler in this part, because we need to know many different concepts which are provided by the Linux kernel, but will see it in the chapter about the memory management in the Linux kernel.

Back to start_kernel

There are many different function calls after the `early_trap_pf_init` in the `setup_arch` function from different kernel subsystems, but there are no one interrupts and exceptions handling related. So, we have to go back where we came from - `start_kernel` function from the init/main.c. The first things after the `setup_arch` is the `trap_init` function from the arch/x86/kernel/traps.c. This function makes initialization of the remaining exceptions handlers (remember that we already setup 3 handlers for the `#DB` - debug exception, `#BP` - breakpoint exception and `#PF` - page fault exception). The `trap_init` function starts from the check of the Extended Industry Standard Architecture:

```
#ifdef CONFIG_EISA
        void __iomem *p = early_ioremap(0x0FFFD9, 4);

        if (readl(p) == 'E' + ('I'<<8) + ('S'<<16) + ('A'<<24))
                EISA_bus = 1;
        early_iounmap(p, 4);
#endif
```

Note that it depends on the `CONFIG_EISA` kernel configuration parameter which represents `EISA` support. Here we use `early_ioremap` function to map `I/O` memory on the page tables. We use `readl` function to read first `4` bytes from the mapped region and if they are equal to `EISA` string we set `EISA_bus` to one. In the end we just unmap previously mapped region. More about `early_ioremap` you can read in the part which describes Fix-Mapped Addresses and ioremap.

After this we start to fill the `Interrupt Descriptor Table` with the different interrupt gates. First of all we set `#DE` or `Divide Error` and `#NMI` or `Non-maskable Interrupt` :

```
set_intr_gate(X86_TRAP_DE, divide_error);
set_intr_gate_ist(X86_TRAP_NMI, &nmi, NMI_STACK);
```

We use `set_intr_gate` macro to set the interrupt gate for the `#DE` exception and `set_intr_gate_ist` for the `#NMI` . You can remember that we already used these macros when we have set the interrupts gates for the page fault handler, debug handler and etc, you can find explanation of it in the previous part. After this we setup exception gates for the following exceptions:

```
set_system_intr_gate(X86_TRAP_OF, &overflow);
set_intr_gate(X86_TRAP_BR, bounds);
set_intr_gate(X86_TRAP_UD, invalid_op);
set_intr_gate(X86_TRAP_NM, device_not_available);
```

Here we can see:

- `#OF` or `Overflow` exception. This exception indicates that an overflow trap occurred when an special INTO instruction was executed;
- `#BR` or `BOUND Range exceeded` exception. This exception indicates that a `BOUND-range-exceed` fault occurred when a BOUND instruction was executed;
- `#UD` or `Invalid Opcode` exception. Occurs when a processor attempted to execute invalid or reserved opcode, processor attempted to execute instruction with invalid operand(s) and etc;
- `#NM` or `Device Not Available` exception. Occurs when the processor tries to execute `x87 FPU` floating point instruction while `EM` flag in the control register `cr0` was set.

In the next step we set the interrupt gate for the `#DF` or `Double fault` exception:

```
set_intr_gate_ist(X86_TRAP_DF, &double_fault, DOUBLEFAULT_STACK);
```

This exception occurs when processor detected a second exception while calling an exception handler for a prior exception. In usual way when the processor detects another exception while trying to call an exception handler, the two exceptions can be handled serially. If the processor cannot handle them serially, it signals the double-fault or `#DF` exception.

The following set of the interrupt gates is:

```
set_intr_gate(X86_TRAP_OLD_MF, &coprocessor_segment_overrun);
set_intr_gate(X86_TRAP_TS, &invalid_TSS);
set_intr_gate(X86_TRAP_NP, &segment_not_present);
set_intr_gate_ist(X86_TRAP_SS, &stack_segment, STACKFAULT_STACK);
set_intr_gate(X86_TRAP_GP, &general_protection);
set_intr_gate(X86_TRAP_SPURIOUS, &spurious_interrupt_bug);
set_intr_gate(X86_TRAP_MF, &coprocessor_error);
set_intr_gate(X86_TRAP_AC, &alignment_check);
```

Here we can see setup for the following exception handlers:

- `#CSO` or `Coprocessor Segment Overrun` - this exception indicates that math coprocessor of an old processor detected a page or segment violation. Modern processors do not generate this exception
- `#TS` or `Invalid TSS` exception - indicates that there was an error related to the Task State Segment.
- `#NP` or `Segment Not Present` exception indicates that the `present flag` of a segment or gate descriptor is clear during attempt to load one of `cs` , `ds` , `es` , `fs` , or `gs` register.
- `#SS` or `Stack Fault` exception indicates one of the stack related conditions was detected, for example a not-present stack segment is detected when attempting to load the `ss` register.
- `#GP` or `General Protection` exception indicates that the processor detected one of a class of protection violations called

general-protection violations. There are many different conditions that can cause general-protection exception. For example loading the `ss`, `ds`, `es`, `fs`, or `gs` register with a segment selector for a system segment, writing to a code segment or a read-only data segment, referencing an entry in the `Interrupt Descriptor Table` (following an interrupt or exception) that is not an interrupt, trap, or task gate and many many more.
- `Spurious Interrupt` - a hardware interrupt that is unwanted.
- `#MF` or `x87 FPU Floating-Point Error` exception caused when the x87 FPU has detected a floating point error.
- `#AC` or `Alignment Check` exception Indicates that the processor detected an unaligned memory operand when alignment checking was enabled.

After that we setup this exception gates, we can see setup of the `Machine-Check` exception:

```
#ifdef CONFIG_X86_MCE
    set_intr_gate_ist(X86_TRAP_MC, &machine_check, MCE_STACK);
#endif
```

Note that it depends on the `CONFIG_X86_MCE` kernel configuration option and indicates that the processor detected an internal machine error or a bus error, or that an external agent detected a bus error. The next exception gate is for the SIMD Floating-Point exception:

```
set_intr_gate(X86_TRAP_XF, &simd_coprocessor_error);
```

which indicates the processor has detected an `SSE` or `SSE2` or `SSE3` SIMD floating-point exception. There are six classes of numeric exception conditions that can occur while executing an SIMD floating-point instruction:

- Invalid operation
- Divide-by-zero
- Denormal operand
- Numeric overflow
- Numeric underflow
- Inexact result (Precision)

In the next step we fill the `used_vectors` array which defined in the arch/x86/include/asm/desc.h header file and represents `bitmap`:

```
DECLARE_BITMAP(used_vectors, NR_VECTORS);
```

of the first `32` interrupts (more about bitmaps in the Linux kernel you can read in the part which describes cpumasks and bitmaps)

```
for (i = 0; i < FIRST_EXTERNAL_VECTOR; i++)
    set_bit(i, used_vectors)
```

where `FIRST_EXTERNAL_VECTOR` is:

```
#define FIRST_EXTERNAL_VECTOR           0x20
```

After this we setup the interrupt gate for the `ia32_syscall` and add `0x80` to the `used_vectors` bitmap:

```
#ifdef CONFIG_IA32_EMULATION
        set_system_intr_gate(IA32_SYSCALL_VECTOR, ia32_syscall);
        set_bit(IA32_SYSCALL_VECTOR, used_vectors);
#endif
```

There is `CONFIG_IA32_EMULATION` kernel configuration option on `x86_64` Linux kernels. This option provides ability to execute 32-bit processes in compatibility-mode. In the next parts we will see how it works, in the meantime we need only to know that there is yet another interrupt gate in the `IDT` with the vector number `0x80`. In the next step we maps `IDT` to the fixmap area:

```
__set_fixmap(FIX_RO_IDT, __pa_symbol(idt_table), PAGE_KERNEL_RO);
idt_descr.address = fix_to_virt(FIX_RO_IDT);
```

and write its address to the `idt_descr.address` (more about fix-mapped addresses you can read in the second part of the Linux kernel memory management chapter). After this we can see the call of the `cpu_init` function that defined in the arch/x86/kernel/cpu/common.c. This function makes initialization of the all `per-cpu` state. In the beginning of the `cpu_init` we do the following things: First of all we wait while current cpu is initialized and than we call the `cr4_init_shadow` function which stores shadow copy of the `cr4` control register for the current cpu and load CPU microcode if need with the following function calls:

```
wait_for_master_cpu(cpu);
cr4_init_shadow();
load_ucode_ap();
```

Next we get the `Task State Segment` for the current cpu and `orig_ist` structure which represents origin `Interrupt Stack Table` values with the:

```
t = &per_cpu(cpu_tss, cpu);
oist = &per_cpu(orig_ist, cpu);
```

As we got values of the `Task State Segment` and `Interrupt Stack Table` for the current processor, we clear following bits in the `cr4` control register:

```
cr4_clear_bits(X86_CR4_VME|X86_CR4_PVI|X86_CR4_TSD|X86_CR4_DE);
```

with this we disable `vm86` extension, virtual interrupts, timestamp (RDTSC can only be executed with the highest privilege) and debug extension. After this we reload the `Global Descriptor Table` and `Interrupt Descriptor table` with the:

```
    switch_to_new_gdt(cpu);
    loadsegment(fs, 0);
    load_current_idt();
```

After this we setup array of the Thread-Local Storage Descriptors, configure NX and load CPU microcode. Now is time to setup and load `per-cpu` Task State Segments. We are going in a loop through the all exception stack which is `N_EXCEPTION_STACKS` or `4` and fill it with `Interrupt Stack Tables`:

```
    if (!oist->ist[0]) {
        char *estacks = per_cpu(exception_stacks, cpu);

        for (v = 0; v < N_EXCEPTION_STACKS; v++) {
            estacks += exception_stack_sizes[v];
            oist->ist[v] = t->x86_tss.ist[v] =
                    (unsigned long)estacks;
            if (v == DEBUG_STACK-1)
                per_cpu(debug_stack_addr, cpu) = (unsigned long)estacks;
        }
    }
```

As we have filled `Task State Segments` with the `Interrupt Stack Tables` we can set `TSS` descriptor for the current processor and load it with the:

```
set_tss_desc(cpu, t);
load_TR_desc();
```

where `set_tss_desc` macro from the [arch/x86/include/asm/desc.h](https://github.com/torvalds/linux/blob/master/arch/x86/include/asm/desc.h) writes given descriptor to the `Global Descriptor Table` of the given processor:

```C
#define set_tss_desc(cpu, addr) __set_tss_desc(cpu, GDT_ENTRY_TSS, addr)
static inline void __set_tss_desc(unsigned cpu, unsigned int entry, void *addr)
{
        struct desc_struct *d = get_cpu_gdt_table(cpu);
        tss_desc tss;
        set_tssldt_descriptor(&tss, (unsigned long)addr, DESC_TSS,
                              IO_BITMAP_OFFSET + IO_BITMAP_BYTES +
                              sizeof(unsigned long) - 1);
        write_gdt_entry(d, entry, &tss, DESC_TSS);
}
```

and `load_TR_desc` macro expands to the `ltr` or `Load Task Register` instruction:

```C
#define load_TR_desc()                          native_load_tr_desc()
static inline void native_load_tr_desc(void)
{
        asm volatile("ltr %w0"::"q" (GDT_ENTRY_TSS*8));
}
```

In the end of the `trap_init` function we can see the following code:

```C
set_intr_gate_ist(X86_TRAP_DB, &debug, DEBUG_STACK);
set_system_intr_gate_ist(X86_TRAP_BP, &int3, DEBUG_STACK);
...
...
...
#ifdef CONFIG_X86_64
        memcpy(&nmi_idt_table, &idt_table, IDT_ENTRIES * 16);
        set_nmi_gate(X86_TRAP_DB, &debug);
        set_nmi_gate(X86_TRAP_BP, &int3);
#endif
```

Here we copy `idt_table` to the `nmi_dit_table` and setup exception handlers for the `#DB` or `Debug exception` and `#BR` or `Breakpoint exception`. You can remember that we already set these interrupt gates in the previous [part](https://0xax.gitbooks.io/linux-insides/content/interrupts/interrupts-3.html), so why do we need to setup it again? We setup it again because when we initialized it before in the `early_trap_init` function, the `Task State Segment` was not ready yet, but now it is ready after the call of the `cpu_init` function.

That's all. Soon we will consider all handlers of these interrupts/exceptions.

Conclusion

It is the end of the fourth part about interrupts and interrupt handling in the Linux kernel. We saw the initialization of the [Task State Segment](https://en.wikipedia.org/wiki/Task_state_segment) in this part and initialization of the different interrupt handlers as `Divide Error`, `Page Fault` exception and etc. You can note that we saw just initialization stuff, and will dive into details about handlers for these exceptions. In the next part we will start to do it.

If you have any questions or suggestions write me a comment or ping me at [twitter](https://twitter.com/0xAX).

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to [linux-insides](https://github.com/0xAX/linux-insides).

Links

- page fault
- Interrupt Descriptor Table
- Tracing
- cr2
- RCU
- this*cpu** operations
- kmemcheck
- prefetchw
- 3DNow
- CPU caches
- VFS
- Linux kernel memory management
- Fix-Mapped Addresses and ioremap
- Extended Industry Standard Architecture
- INT isntruction
- INTO
- BOUND
- opcode
- control register
- x87 FPU
- MCE exception
- SIMD
- cpumasks and bitmaps
- NX
- Task State Segment
- Previous part

Interrupts and Interrupt Handling. Part 5.

Implementation of exception handlers

This is the fifth part about an interrupts and exceptions handling in the Linux kernel and in the previous part we stopped on the setting of interrupt gates to the Interrupt descriptor Table. We did it in the `trap_init` function from the arch/x86/kernel/traps.c source code file. We saw only setting of these interrupt gates in the previous part and in the current part we will see implementation of the exception handlers for these gates. The preparation before an exception handler will be executed is in the arch/x86/entry/entry_64.S assembly file and occurs in the idtentry macro that defines exceptions entry points:

```
idtentry divide_error                    do_divide_error                  has_error_code=0
idtentry overflow                        do_overflow                      has_error_code=0
idtentry invalid_op                        do_invalid_op                  has_error_code=0
idtentry bounds                            do_bounds                      has_error_code=0
idtentry device_not_available            do_device_not_available          has_error_code=0
idtentry coprocessor_segment_overrun    do_coprocessor_segment_overrun has_error_code=0
idtentry invalid_TSS                    do_invalid_TSS                  has_error_code=1
idtentry segment_not_present            do_segment_not_present          has_error_code=1
idtentry spurious_interrupt_bug            do_spurious_interrupt_bug      has_error_code=0
idtentry coprocessor_error                do_coprocessor_error            has_error_code=0
idtentry alignment_check                do_alignment_check              has_error_code=1
idtentry simd_coprocessor_error            do_simd_coprocessor_error      has_error_code=0
```

The `idtentry` macro does following preparation before an actual exception handler (`do_divide_error` for the `divide_error` , `do_overflow` for the `overflow` and etc.) will get control. In another words the `idtentry` macro allocates place for the registers (pt_regs structure) on the stack, pushes dummy error code for the stack consistency if an interrupt/exception has no error code, checks the segment selector in the `cs` segment register and switches depends on the previous state(userspace or kernelspace). After all of these preparations it makes a call of an actual interrupt/exception handler:

```
.macro idtentry sym do_sym has_error_code:req paranoid=0 shift_ist=-1
ENTRY(\sym)
    ...
    ...
    ...
    call    \do_sym
    ...
    ...
    ...
END(\sym)
.endm
```

After an exception handler will finish its work, the `idtentry` macro restores stack and general purpose registers of an interrupted task and executes iret instruction:

```
ENTRY(paranoid_exit)
    ...
    ...
    ...
    RESTORE_EXTRA_REGS
    RESTORE_C_REGS
    REMOVE_PT_GPREGS_FROM_STACK 8
    INTERRUPT_RETURN
END(paranoid_exit)
```

where `INTERRUPT_RETURN` is:

```
#define INTERRUPT_RETURN    jmp native_iret
...
ENTRY(native_iret)
.global native_irq_return_iret
native_irq_return_iret:
iretq
```

More about the `idtentry` macro you can read in the third part of the https://0xax.gitbooks.io/linux-insides/content/Interrupts/linux-interrupts-3.html chapter. Ok, now we saw the preparation before an exception handler will be executed and now time to look on the handlers. First of all let's look on the following handlers:

- divide_error
- overflow
- invalid_op
- coprocessor_segment_overrun
- invalid_TSS
- segment_not_present
- stack_segment
- alignment_check

All these handlers defined in the arch/x86/kernel/traps.c source code file with the `DO_ERROR` macro:

```
DO_ERROR(X86_TRAP_DE,     SIGFPE,  "divide error",                divide_error)
DO_ERROR(X86_TRAP_OF,     SIGSEGV, "overflow",                    overflow)
DO_ERROR(X86_TRAP_UD,     SIGILL,  "invalid opcode",              invalid_op)
DO_ERROR(X86_TRAP_OLD_MF, SIGFPE,  "coprocessor segment overrun", coprocessor_segment_overrun)
DO_ERROR(X86_TRAP_TS,     SIGSEGV, "invalid TSS",                 invalid_TSS)
DO_ERROR(X86_TRAP_NP,     SIGBUS,  "segment not present",         segment_not_present)
DO_ERROR(X86_TRAP_SS,     SIGBUS,  "stack segment",               stack_segment)
DO_ERROR(X86_TRAP_AC,     SIGBUS,  "alignment check",             alignment_check)
```

As we can see the `DO_ERROR` macro takes 4 parameters:

- Vector number of an interrupt;
- Signal number which will be sent to the interrupted process;
- String which describes an exception;
- Exception handler entry point.

This macro defined in the same source code file and expands to the function with the `do_handler` name:

```
#define DO_ERROR(trapnr, signr, str, name)                              \
dotraplinkage void do_##name(struct pt_regs *regs, long error_code)     \
{                                                                       \
        do_error_trap(regs, error_code, str, trapnr, signr);            \
}
```

Note on the `##` tokens. This is special feature - GCC macro Concatenation which concatenates two given strings. For example, first `DO_ERROR` in our example will expands to the:

```
dotraplinkage void do_divide_error(struct pt_regs *regs, long error_code)     \
{
    ...
}
```

We can see that all functions which are generated by the `DO_ERROR` macro just make a call of the `do_error_trap` function from the arch/x86/kernel/traps.c. Let's look on implementation of the `do_error_trap` function.

Trap handlers

The `do_error_trap` function starts and ends from the two following functions:

```
enum ctx_state prev_state = exception_enter();
...
...
...
exception_exit(prev_state);
```

from the include/linux/context_tracking.h. The context tracking in the Linux kernel subsystem which provide kernel boundaries probes to keep track of the transitions between level contexts with two basic initial contexts: `user` or `kernel` . The `exception_enter` function checks that context tracking is enabled. After this if it is enabled, the `exception_enter` reads previous context and compares it with the `CONTEXT_KERNEL` . If the previous context is `user` , we call `context_tracking_exit` function from the kernel/context_tracking.c which inform the context tracking subsystem that a processor is exiting user mode and entering the kernel mode:

```
if (!context_tracking_is_enabled())
    return 0;

prev_ctx = this_cpu_read(context_tracking.state);
if (prev_ctx != CONTEXT_KERNEL)
    context_tracking_exit(prev_ctx);

return prev_ctx;
```

If previous context is non `user` , we just return it. The `pre_ctx` has `enum ctx_state` type which defined in the include/linux/context_tracking_state.h and looks as:

```
enum ctx_state {
    CONTEXT_KERNEL = 0,
    CONTEXT_USER,
    CONTEXT_GUEST,
} state;
```

The second function is `exception_exit` defined in the same include/linux/context_tracking.h file and checks that context tracking is enabled and call the `contert_tracking_enter` function if the previous context was `user` :

```
static inline void exception_exit(enum ctx_state prev_ctx)
{
    if (context_tracking_is_enabled()) {
        if (prev_ctx != CONTEXT_KERNEL)
            context_tracking_enter(prev_ctx);
    }
}
```

The `context_tracking_enter` function informs the context tracking subsystem that a processor is going to enter to the user mode from the kernel mode. We can see the following code between the `exception_enter` and `exception_exit` :

```
if (notify_die(DIE_TRAP, str, regs, error_code, trapnr, signr) !=
        NOTIFY_STOP) {
    conditional_sti(regs);
    do_trap(trapnr, signr, str, regs, error_code,
        fill_trap_info(regs, signr, trapnr, &info));
}
```

First of all it calls the `notify_die` function which defined in the kernel/notifier.c. To get notified for kernel panic, kernel oops, Non-Maskable Interrupt or other events the caller needs to insert itself in the `notify_die` chain and the `notify_die` function does it. The Linux kernel has special mechanism that allows kernel to ask when something happens and this mechanism called `notifiers` or `notifier chains` . This mechanism used for example for the `USB` hotplug events (look on the drivers/usb/core/notify.c), for the memory hotplug (look on the include/linux/memory.h, the `hotplug_memory_notifier` macro and etc...), system reboots and etc. A notifier chain is thus a simple, singly-linked list. When a Linux kernel subsystem wants to be notified of specific events, it fills out a special `notifier_block` structure and passes it to the `notifier_chain_register` function. An event can be sent with the call of the `notifier_call_chain` function. First of all the `notify_die` function fills `die_args` structure with the trap number, trap string, registers and other values:

```
struct die_args args = {
       .regs   = regs,
       .str    = str,
       .err    = err,
       .trapnr = trap,
       .signr  = sig,
}
```

and returns the result of the `atomic_notifier_call_chain` function with the `die_chain` :

```
static ATOMIC_NOTIFIER_HEAD(die_chain);
return atomic_notifier_call_chain(&die_chain, val, &args);
```

which just expands to the `atomic_notifier_head` structure that contains lock and `notifier_block` :

```
struct atomic_notifier_head {
        spinlock_t lock;
        struct notifier_block __rcu *head;
};
```

The `atomic_notifier_call_chain` function calls each function in a notifier chain in turn and returns the value of the last notifier function called. If the `notify_die` in the `do_error_trap` does not return `NOTIFY_STOP` we execute `conditional_sti` function from the arch/x86/kernel/traps.c that checks the value of the interrupt flag and enables interrupt depends on it:

```
static inline void conditional_sti(struct pt_regs *regs)
{
        if (regs->flags & X86_EFLAGS_IF)
                local_irq_enable();
}
```

more about `local_irq_enable` macro you can read in the second part of this chapter. The next and last call in the `do_error_trap` is the `do_trap` function. First of all the `do_trap` function defined the `tsk` variable which has `task_struct` type and represents the current interrupted process. After the definition of the `tsk` , we can see the call of the `do_trap_no_signal` function:

```
struct task_struct *tsk = current;

if (!do_trap_no_signal(tsk, trapnr, str, regs, error_code))
    return;
```

The `do_trap_no_signal` function makes two checks:

- Did we come from the Virtual 8086 mode;
- Did we come from the kernelspace.

```
if (v8086_mode(regs)) {
    ...
}

if (!user_mode(regs)) {
    ...
}

return -1;
```

We will not consider first case because the long mode does not support the Virtual 8086 mode. In the second case we invoke `fixup_exception` function which will try to recover a fault and `die` if we can't:

```
if (!fixup_exception(regs)) {
    tsk->thread.error_code = error_code;
    tsk->thread.trap_nr = trapnr;
    die(str, regs, error_code);
}
```

The `die` function defined in the arch/x86/kernel/dumpstack.c source code file, prints useful information about stack, registers, kernel modules and caused kernel oops. If we came from the userspace the `do_trap_no_signal` function will return `-1` and the execution of the `do_trap` function will continue. If we passed through the `do_trap_no_signal` function and did not exit from the `do_trap` after this, it means that previous context was - `user` . Most exceptions caused by the processor are interpreted by Linux as error conditions, for example division by zero, invalid opcode and etc. When an exception occurs the Linux kernel sends a signal to the interrupted process that caused the exception to notify it of an incorrect condition. So, in the `do_trap` function we need to send a signal with the given number (`SIGFPE` for the divide error, `SIGILL` for the overflow exception and etc...). First of all we save error code and vector number in the current interrupts process with the filling `thread.error_code` and `thread_trap_nr` :

```
tsk->thread.error_code = error_code;
tsk->thread.trap_nr = trapnr;
```

After this we make a check do we need to print information about unhandled signals for the interrupted process. We check that `show_unhandled_signals` variable is set, that `unhandled_signal` function from the kernel/signal.c will return unhandled signal(s) and printk rate limit:

```
#ifdef CONFIG_X86_64
    if (show_unhandled_signals && unhandled_signal(tsk, signr) &&
        printk_ratelimit()) {
        pr_info("%s[%d] trap %s ip:%lx sp:%lx error:%lx",
            tsk->comm, tsk->pid, str,
            regs->ip, regs->sp, error_code);
        print_vma_addr(" in ", regs->ip);
        pr_cont("\n");
    }
#endif
```

And send a given signal to interrupted process:

```
force_sig_info(signr, info ?: SEND_SIG_PRIV, tsk);
```

This is the end of the `do_trap` . We just saw generic implementation for eight different exceptions which are defined with the `DO_ERROR` macro. Now let's look on another exception handlers.

Double fault

The next exception is `#DF` or `Double fault`. This exception occurs when the processor detected a second exception while calling an exception handler for a prior exception. We set the trap gate for this exception in the previous part:

```
set_intr_gate_ist(X86_TRAP_DF, &double_fault, DOUBLEFAULT_STACK);
```

Note that this exception runs on the `DOUBLEFAULT_STACK` Interrupt Stack Table which has index - `1`:

```
#define DOUBLEFAULT_STACK 1
```

The `double_fault` is handler for this exception and defined in the arch/x86/kernel/traps.c. The `double_fault` handler starts from the definition of two variables: string that describes exception and interrupted process, as other exception handlers:

```
static const char str[] = "double fault";
struct task_struct *tsk = current;
```

The handler of the double fault exception split on two parts. The first part is the check which checks that a fault is a `non-IST` fault on the `espfix64` stack. Actually the `iret` instruction restores only the bottom `16` bits when returning to a `16` bit segment. The `espfix` feature solves this problem. So if the `non-IST` fault on the espfix64 stack we modify the stack to make it look like `General Protection Fault`:

```
struct pt_regs *normal_regs = task_pt_regs(current);

memmove(&normal_regs->ip, (void *)regs->sp, 5*8);
ormal_regs->orig_ax = 0;
regs->ip = (unsigned long)general_protection;
regs->sp = (unsigned long)&normal_regs->orig_ax;
return;
```

In the second case we do almost the same that we did in the previous exception handlers. The first is the call of the `ist_enter` function that discards previous context, `user` in our case:

```
ist_enter(regs);
```

And after this we fill the interrupted process with the vector number of the `Double fault` exception and error code as we did it in the previous handlers:

```
tsk->thread.error_code = error_code;
tsk->thread.trap_nr = X86_TRAP_DF;
```

Next we print useful information about the double fault (PID number, registers content):

```
#ifdef CONFIG_DOUBLEFAULT
    df_debug(regs, error_code);
#endif
```

And die:

```
    for (;;)
        die(str, regs, error_code);
```

That's all.

Device not available exception handler

The next exception is the `#NM` or `Device not available`. The `Device not available` exception can occur depending on these things:

- The processor executed an x87 FPU floating-point instruction while the EM flag in control register `cr0` was set;
- The processor executed a `wait` or `fwait` instruction while the `MP` and `TS` flags of register `cr0` were set;
- The processor executed an x87 FPU, MMX or SSE instruction while the `TS` flag in control register `cr0` was set and the `EM` flag is clear.

The handler of the `Device not available` exception is the `do_device_not_available` function and it defined in the arch/x86/kernel/traps.c source code file too. It starts and ends from the getting of the previous context, as other traps which we saw in the beginning of this part:

```
enum ctx_state prev_state;
prev_state = exception_enter();
...
...
...
exception_exit(prev_state);
```

In the next step we check that `FPU` is not eager:

```
BUG_ON(use_eager_fpu());
```

When we switch into a task or interrupt we may avoid loading the `FPU` state. If a task will use it, we catch `Device not Available exception` exception. If we loading the `FPU` state during task switching, the `FPU` is eager. In the next step we check `cr0` control register on the `EM` flag which can show us is `x87` floating point unit present (flag clear) or not (flag set):

```
#ifdef CONFIG_MATH_EMULATION
    if (read_cr0() & X86_CR0_EM) {
        struct math_emu_info info = { };

        conditional_sti(regs);

        info.regs = regs;
        math_emulate(&info);
        exception_exit(prev_state);
        return;
    }
#endif
```

If the `x87` floating point unit not presented, we enable interrupts with the `conditional_sti`, fill the `math_emu_info` (defined in the arch/x86/include/asm/math_emu.h) structure with the registers of an interrupt task and call `math_emulate` function from the arch/x86/math-emu/fpu_entry.c. As you can understand from function's name, it emulates `X87 FPU` unit (more about the `x87` we will know in the special chapter). In other way, if `X86_CR0_EM` flag is clear which means that `x87 FPU` unit is presented, we call the `fpu__restore` function from the arch/x86/kernel/fpu/core.c which copies the `FPU` registers from the `fpustate` to the live hardware registers. After this `FPU` instructions can be used:

```
fpu__restore(&current->thread.fpu);
```

General protection fault exception handler

The next exception is the `#GP` or `General protection fault`. This exception occurs when the processor detected one of a class of protection violations called `general-protection violations`. It can be:

- Exceeding the segment limit when accessing the `cs`, `ds`, `es`, `fs` or `gs` segments;
- Loading the `ss`, `ds`, `es`, `fs` or `gs` register with a segment selector for a system segment.;
- Violating any of the privilege rules;
- and other...

The exception handler for this exception is the `do_general_protection` from the arch/x86/kernel/traps.c. The `do_general_protection` function starts and ends as other exception handlers from the getting of the previous context:

```
prev_state = exception_enter();
...
exception_exit(prev_state);
```

After this we enable interrupts if they were disabled and check that we came from the Virtual 8086 mode:

```
conditional_sti(regs);

if (v8086_mode(regs)) {
    local_irq_enable();
    handle_vm86_fault((struct kernel_vm86_regs *) regs, error_code);
    goto exit;
}
```

As long mode does not support this mode, we will not consider exception handling for this case. In the next step check that previous mode was kernel mode and try to fix the trap. If we can't fix the current general protection fault exception we fill the interrupted process with the vector number and error code of the exception and add it to the `notify_die` chain:

```
if (!user_mode(regs)) {
    if (fixup_exception(regs))
        goto exit;

    tsk->thread.error_code = error_code;
    tsk->thread.trap_nr = X86_TRAP_GP;
    if (notify_die(DIE_GPF, "general protection fault", regs, error_code,
               X86_TRAP_GP, SIGSEGV) != NOTIFY_STOP)
        die("general protection fault", regs, error_code);
    goto exit;
}
```

If we can fix exception we go to the `exit` label which exits from exception state:

```
exit:
    exception_exit(prev_state);
```

If we came from user mode we send `SIGSEGV` signal to the interrupted process from user mode as we did it in the `do_trap` function:

```
if (show_unhandled_signals && unhandled_signal(tsk, SIGSEGV) &&
        printk_ratelimit()) {
    pr_info("%s[%d] general protection ip:%lx sp:%lx error:%lx",
        tsk->comm, task_pid_nr(tsk),
        regs->ip, regs->sp, error_code);
    print_vma_addr(" in ", regs->ip);
    pr_cont("\n");
}

force_sig_info(SIGSEGV, SEND_SIG_PRIV, tsk);
```

That's all.

Conclusion

It is the end of the fifth part of the Interrupts and Interrupt Handling chapter and we saw implementation of some interrupt handlers in this part. In the next part we will continue to dive into interrupt and exception handlers and will see handler for the Non-Maskable Interrupts, handling of the math coprocessor and SIMD coprocessor exceptions and many many more.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- Interrupt descriptor Table
- iret instruction
- GCC macro Concatenation
- kernel panic
- kernel oops
- Non-Maskable Interrupt
- hotplug
- interrupt flag
- long mode
- signal
- printk
- coprocessor
- SIMD
- Interrupt Stack Table
- PID
- x87 FPU
- control register
- MMX
- Previous part

Interrupts and Interrupt Handling. Part 6.

Non-maskable interrupt handler

It is sixth part of the Interrupts and Interrupt Handling in the Linux kernel chapter and in the previous part we saw implementation of some exception handlers for the General Protection Fault exception, divide exception, invalid opcode exceptions and etc. As I wrote in the previous part we will see implementations of the rest exceptions in this part. We will see implementation of the following handlers:

- Non-Maskable interrupt;
- BOUND Range Exceeded Exception;
- Coprocessor exception;
- SIMD coprocessor exception.

in this part. So, let's start.

Non-Maskable interrupt handling

A Non-Maskable interrupt is a hardware interrupt that cannot be ignored by standard masking techniques. In a general way, a non-maskable interrupt can be generated in either of two ways:

- External hardware asserts the non-maskable interrupt pin on the CPU.
- The processor receives a message on the system bus or the APIC serial bus with a delivery mode `NMI` .

When the processor receives a `NMI` from one of these sources, the processor handles it immediately by calling the `NMI` handler pointed to by interrupt vector which has number `2` (see table in the first part). We already filled the Interrupt Descriptor Table with the vector number, address of the `nmi` interrupt handler and `NMI_STACK` Interrupt Stack Table entry:

```
set_intr_gate_ist(X86_TRAP_NMI, &nmi, NMI_STACK);
```

in the `trap_init` function which defined in the arch/x86/kernel/traps.c source code file. In the previous parts we saw that entry points of the all interrupt handlers are defined with the:

```
.macro idtentry sym do_sym has_error_code:req paranoid=0 shift_ist=-1
ENTRY(\sym)
...
...
...
END(\sym)
.endm
```

macro from the arch/x86/entry/entry_64.S assembly source code file. But the handler of the `Non-Maskable` interrupts is not defined with this macro. It has own entry point:

```
ENTRY(nmi)
...
...
...
END(nmi)
```

in the same arch/x86/entry/entry_64.S assembly file. Lets dive into it and will try to understand how `Non-Maskable` interrupt handler works. The `nmi` handlers starts from the call of the:

```
PARAVIRT_ADJUST_EXCEPTION_FRAME
```

macro but we will not dive into details about it in this part, because this macro related to the Paravirtualization stuff which we will see in another chapter. After this save the content of the `rdx` register on the stack:

```
pushq    %rdx
```

And allocated check that `cs` was not the kernel segment when an non-maskable interrupt occurs:

```
cmpl    $__KERNEL_CS, 16(%rsp)
jne    first_nmi
```

The `__KERNEL_CS` macro defined in the arch/x86/include/asm/segment.h and represented second descriptor in the Global Descriptor Table:

```
#define GDT_ENTRY_KERNEL_CS    2
#define __KERNEL_CS    (GDT_ENTRY_KERNEL_CS*8)
```

more about `GDT` you can read in the second part of the Linux kernel booting process chapter. If `cs` is not kernel segment, it means that it is not nested `NMI` and we jump on the `first_nmi` label. Let's consider this case. First of all we put address of the current stack pointer to the `rdx` and pushes `1` to the stack in the `first_nmi` label:

```
first_nmi:
    movq    (%rsp), %rdx
    pushq    $1
```

Why do we push `1` on the stack? As the comment says: `We allow breakpoints in NMIs`. On the x86_64, like other architectures, the CPU will not execute another `NMI` until the first `NMI` is completed. A `NMI` interrupt finished with the iret instruction like other interrupts and exceptions do it. If the `NMI` handler triggers either a page fault or breakpoint or another exception which are use `iret` instruction too. If this happens while in `NMI` context, the CPU will leave `NMI` context and a new `NMI` may come in. The `iret` used to return from those exceptions will re-enable `NMIs` and we will get nested non-maskable interrupts. The problem the `NMI` handler will not return to the state that it was, when the exception triggered, but instead it will return to a state that will allow new `NMIs` to preempt the running `NMI` handler. If another `NMI` comes in before the first NMI handler is complete, the new NMI will write all over the preempted `NMIs` stack. We can have nested `NMIs` where the next `NMI` is using the top of the stack of the previous `NMI`. It means that we cannot execute it because a nested non-maskable interrupt will corrupt stack of a previous non-maskable interrupt. That's why we have allocated space on the stack for temporary variable. We will check this variable that it was set when a previous `NMI` is executing and clear if it is not nested `NMI`. We push `1` here to the previously allocated space on the stack to denote that a `non-maskable` interrupt executed currently. Remember that when and `NMI` or another exception occurs we have the following stack frame:

```
+------------------------+
|         SS             |
|         RSP            |
|        RFLAGS          |
|         CS             |
|         RIP            |
+------------------------+
```

and also an error code if an exception has it. So, after all of these manipulations our stack frame will look like this:

```
+------------------------+
|         SS             |
|         RSP            |
```

```
|        RFLAGS          |
|         CS             |
|         RIP            |
|         RDX            |
|          1             |
+------------------------+
```

In the next step we allocate yet another `40` bytes on the stack:

```
subq    $(5*8), %rsp
```

and pushes the copy of the original stack frame after the allocated space:

```
.rept 5
pushq    11*8(%rsp)
.endr
```

with the .rept assembly directive. We need in the copy of the original stack frame. Generally we need in two copies of the interrupt stack. First is `copied` interrupts stack: `saved` stack frame and `copied` stack frame. Now we pushes original stack frame to the `saved` stack frame which locates after the just allocated `40` bytes (`copied` stack frame). This stack frame is used to fixup the `copied` stack frame that a nested NMI may change. The second - `copied` stack frame modified by any nested `NMIs` to let the first `NMI` know that we triggered a second `NMI` and we should repeat the first `NMI` handler. Ok, we have made first copy of the original stack frame, now time to make second copy:

```
addq    $(10*8), %rsp

.rept 5
pushq    -6*8(%rsp)
.endr
subq    $(5*8), %rsp
```

After all of these manipulations our stack frame will be like this:

```
+-------------------------+
| original SS             |
| original Return RSP     |
| original RFLAGS         |
| original CS             |
| original RIP            |
+-------------------------+
| temp storage for rdx    |
+-------------------------+
| NMI executing variable  |
+-------------------------+
| copied SS               |
| copied Return RSP       |
| copied RFLAGS           |
| copied CS               |
| copied RIP              |
+-------------------------+
| Saved SS                |
| Saved Return RSP        |
| Saved RFLAGS            |
| Saved CS                |
| Saved RIP               |
+-------------------------+
```

After this we push dummy error code on the stack as we did it already in the previous exception handlers and allocate space for the general purpose registers on the stack:

```
pushq    $-1
ALLOC_PT_GPREGS_ON_STACK
```

We already saw implementation of the `ALLOC_PT_GREGS_ON_STACK` macro in the third part of the interrupts chapter. This macro defined in the arch/x86/entry/calling.h and yet another allocates `120` bytes on stack for the general purpose registers, from the `rdi` to the `r15` :

```
.macro ALLOC_PT_GPREGS_ON_STACK addskip=0
addq    $-(15*8+\addskip), %rsp
.endm
```

After space allocation for the general registers we can see call of the `paranoid_entry` :

```
call    paranoid_entry
```

We can remember from the previous parts this label. It pushes general purpose registers on the stack, reads `MSR_GS_BASE` Model Specific register and checks its value. If the value of the `MSR_GS_BASE` is negative, we came from the kernel mode and just return from the `paranoid_entry` , in other way it means that we came from the usermode and need to execute `swapgs` instruction which will change user `gs` with the kernel `gs` :

```
ENTRY(paranoid_entry)
    cld
    SAVE_C_REGS 8
    SAVE_EXTRA_REGS 8
    movl    $1, %ebx
    movl    $MSR_GS_BASE, %ecx
    rdmsr
    testl    %edx, %edx
    js    1f
    SWAPGS
    xorl    %ebx, %ebx
1:    ret
END(paranoid_entry)
```

Note that after the `swapgs` instruction we zeroed the `ebx` register. Next time we will check content of this register and if we executed `swapgs` than `ebx` must contain `0` and `1` in other way. In the next step we store value of the `cr2` control register to the `r12` register, because the `NMI` handler can cause `page fault` and corrupt the value of this control register:

```
movq    %cr2, %r12
```

Now time to call actual `NMI` handler. We push the address of the `pt_regs` to the `rdi` , error code to the `rsi` and call the `do_nmi` handler:

```
movq    %rsp, %rdi
movq    $-1, %rsi
call    do_nmi
```

We will back to the `do_nmi` little later in this part, but now let's look what occurs after the `do_nmi` will finish its execution. After the `do_nmi` handler will be finished we check the `cr2` register, because we can got page fault during `do_nmi` performed and if we got it we restore original `cr2` , in other way we jump on the label `1` . After this we test content of the `ebx` register (remember it must contain `0` if we have used `swapgs` instruction and `1` if we didn't use it) and execute `SWAPGS_UNSAFE_STACK` if it contains `1` or jump to the `nmi_restore` label. The `SWAPGS_UNSAFE_STACK` macro just expands to the `swapgs` instruction. In the `nmi_restore` label we restore general purpose registers, clear allocated space on the stack for this registers, clear our temporary variable and exit from the interrupt handler with the `INTERRUPT_RETURN` macro:

```
    movq    %cr2, %rcx
    cmpq    %rcx, %r12
    je    1f
    movq    %r12, %cr2
1:
    testl    %ebx, %ebx
    jnz    nmi_restore
nmi_swapgs:
    SWAPGS_UNSAFE_STACK
nmi_restore:
    RESTORE_EXTRA_REGS
    RESTORE_C_REGS
    /* Pop the extra iret frame at once */
    REMOVE_PT_GPREGS_FROM_STACK 6*8
    /* Clear the NMI executing stack variable */
    movq    $0, 5*8(%rsp)
    INTERRUPT_RETURN
```

where `INTERRUPT_RETURN` is defined in the arch/x86/include/irqflags.h and just expands to the `iret` instruction. That's all.

Now let's consider case when another `NMI` interrupt occurred when previous `NMI` interrupt didn't finish its execution. You can remember from the beginning of this part that we've made a check that we came from userspace and jump on the `first_nmi` in this case:

```
cmpl    $__KERNEL_CS, 16(%rsp)
jne    first_nmi
```

Note that in this case it is first `NMI` every time, because if the first `NMI` catched page fault, breakpoint or another exception it will be executed in the kernel mode. If we didn't come from userspace, first of all we test our temporary variable:

```
cmpl    $1, -8(%rsp)
je    nested_nmi
```

and if it is set to `1` we jump to the `nested_nmi` label. If it is not `1` , we test the `IST` stack. In the case of nested `NMIs` we check that we are above the `repeat_nmi` . In this case we ignore it, in other way we check that we above than `end_repeat_nmi` and jump on the `nested_nmi_out` label.

Now let's look on the `do_nmi` exception handler. This function defined in the arch/x86/kernel/nmi.c source code file and takes two parameters:

- address of the `pt_regs` ;
- error code.

as all exception handlers. The `do_nmi` starts from the call of the `nmi_nesting_preprocess` function and ends with the call of the `nmi_nesting_postprocess` . The `nmi_nesting_preprocess` function checks that we likely do not work with the debug stack and if we on the debug stack set the `update_debug_stack` per-cpu variable to `1` and call the `debug_stack_set_zero` function from the arch/x86/kernel/cpu/common.c. This function increases the `debug_stack_use_ctr` per-cpu variable and loads new `Interrupt Descriptor Table` :

```
static inline void nmi_nesting_preprocess(struct pt_regs *regs)
{
        if (unlikely(is_debug_stack(regs->sp))) {
                debug_stack_set_zero();
                this_cpu_write(update_debug_stack, 1);
        }
}
```

The `nmi_nesting_postprocess` function checks the `update_debug_stack` per-cpu variable which we set in the `nmi_nesting_preprocess` and resets debug stack or in another words it loads origin `Interrupt Descriptor Table`. After the call of the `nmi_nesting_preprocess` function, we can see the call of the `nmi_enter` in the `do_nmi`. The `nmi_enter` increases `lockdep_recursion` field of the interrupted process, update preempt counter and informs the RCU subsystem about `NMI`. There is also `nmi_exit` function that does the same stuff as `nmi_enter`, but vice-versa. After the `nmi_enter` we increase `__nmi_count` in the `irq_stat` structure and call the `default_do_nmi` function. First of all in the `default_do_nmi` we check the address of the previous nmi and update address of the last nmi to the actual:

```
if (regs->ip == __this_cpu_read(last_nmi_rip))
    b2b = true;
else
    __this_cpu_write(swallow_nmi, false);

__this_cpu_write(last_nmi_rip, regs->ip);
```

After this first of all we need to handle CPU-specific `NMIs`:

```
handled = nmi_handle(NMI_LOCAL, regs, b2b);
__this_cpu_add(nmi_stats.normal, handled);
```

And then non-specific `NMIs` depends on its reason:

```
reason = x86_platform.get_nmi_reason();
if (reason & NMI_REASON_MASK) {
    if (reason & NMI_REASON_SERR)
        pci_serr_error(reason, regs);
    else if (reason & NMI_REASON_IOCHK)
        io_check_error(reason, regs);

    __this_cpu_add(nmi_stats.external, 1);
    return;
}
```

That's all.

Range Exceeded Exception

The next exception is the `BOUND` range exceeded exception. The `BOUND` instruction determines if the first operand (array index) is within the bounds of an array specified the second operand (bounds operand). If the index is not within bounds, a `BOUND` range exceeded exception or `#BR` is occurred. The handler of the `#BR` exception is the `do_bounds` function that defined in the arch/x86/kernel/traps.c. The `do_bounds` handler starts with the call of the `exception_enter` function and ends with the call of the `exception_exit`:

```
prev_state = exception_enter();

if (notify_die(DIE_TRAP, "bounds", regs, error_code,
               X86_TRAP_BR, SIGSEGV) == NOTIFY_STOP)
    goto exit;
...
...
...
exception_exit(prev_state);
return;
```

After we have got the state of the previous context, we add the exception to the `notify_die` chain and if it will return `NOTIFY_STOP` we return from the exception. More about notify chains and the `context tracking` functions you can read in the previous part. In the next step we enable interrupts if they were disabled with the `contidional_sti` function that checks `IF` flag and call the `local_irq_enable` depends on its value:

```
conditional_sti(regs);

if (!user_mode(regs))
    die("bounds", regs, error_code);
```

and check that if we didn't came from user mode we send `SIGSEGV` signal with the `die` function. After this we check is MPX enabled or not, and if this feature is disabled we jump on the `exit_trap` label:

```
if (!cpu_feature_enabled(X86_FEATURE_MPX)) {
    goto exit_trap;
}

where we execute `do_trap` function (more about it you can find in the previous part):

```C
exit_trap:
    do_trap(X86_TRAP_BR, SIGSEGV, "bounds", regs, error_code, NULL);
    exception_exit(prev_state);
```

If `MPX` feature is enabled we check the `BNDSTATUS` with the `get_xsave_field_ptr` function and if it is zero, it means that the `MPX` was not responsible for this exception:

```
bndcsr = get_xsave_field_ptr(XSTATE_BNDCSR);
if (!bndcsr)
        goto exit_trap;
```

After all of this, there is still only one way when `MPX` is responsible for this exception. We will not dive into the details about Intel Memory Protection Extensions in this part, but will see it in another chapter.

Coprocessor exception and SIMD exception

The next two exceptions are x87 FPU Floating-Point Error exception or `#MF` and SIMD Floating-Point Exception or `#XF`. The first exception occurs when the `x87 FPU` has detected floating point error. For example divide by zero, numeric overflow and etc. The second exception occurs when the processor has detected SSE/SSE2/SSE3 `SIMD` floating-point exception. It can be the same as for the `x87 FPU`. The handlers for these exceptions are `do_coprocessor_error` and `do_simd_coprocessor_error` are defined in the arch/x86/kernel/traps.c and very similar on each other. They both make a call of the `math_error` function from the same source code file but pass different vector number. The `do_coprocessor_error` passes `X86_TRAP_MF` vector number to the `math_error`:

```
dotraplinkage void do_coprocessor_error(struct pt_regs *regs, long error_code)
{
    enum ctx_state prev_state;

    prev_state = exception_enter();
    math_error(regs, error_code, X86_TRAP_MF);
    exception_exit(prev_state);
}
```

and `do_simd_coprocessor_error` passes `X86_TRAP_XF` to the `math_error` function:
```

```
dotraplinkage void
do_simd_coprocessor_error(struct pt_regs *regs, long error_code)
{
    enum ctx_state prev_state;

    prev_state = exception_enter();
    math_error(regs, error_code, X86_TRAP_XF);
    exception_exit(prev_state);
}
```

First of all the `math_error` function defines current interrupted task, address of its fpu, string which describes an exception, add it to the `notify_die` chain and return from the exception handler if it will return `NOTIFY_STOP` :

```
    struct task_struct *task = current;
    struct fpu *fpu = &task->thread.fpu;
    siginfo_t info;
    char *str = (trapnr == X86_TRAP_MF) ? "fpu exception" :
                        "simd exception";

    if (notify_die(DIE_TRAP, str, regs, error_code, trapnr, SIGFPE) == NOTIFY_STOP)
        return;
```

After this we check that we are from the kernel mode and if yes we will try to fix an exception with the `fixup_exception` function. If we cannot we fill the task with the exception's error code and vector number and die:

```
if (!user_mode(regs)) {
    if (!fixup_exception(regs)) {
        task->thread.error_code = error_code;
        task->thread.trap_nr = trapnr;
        die(str, regs, error_code);
    }
    return;
}
```

If we came from the user mode, we save the `fpu` state, fill the task structure with the vector number of an exception and `siginfo_t` with the number of signal, `errno` , the address where exception occurred and signal code:

```
fpu__save(fpu);

task->thread.trap_nr    = trapnr;
task->thread.error_code = error_code;
info.si_signo        = SIGFPE;
info.si_errno        = 0;
info.si_addr        = (void __user *)uprobe_get_trap_addr(regs);
info.si_code = fpu__exception_code(fpu, trapnr);
```

After this we check the signal code and if it is non-zero we return:

```
if (!info.si_code)
    return;
```

Or send the `SIGFPE` signal in the end:

```
force_sig_info(SIGFPE, &info, task);
```

That's all.

Conclusion

It is the end of the sixth part of the Interrupts and Interrupt Handling chapter and we saw implementation of some exception handlers in this part, like `non-maskable` interrupt, SIMD and x87 FPU floating point exception. Finally we have finsihed with the `trap_init` function in this part and will go ahead in the next part. The next our point is the external interrupts and the `early_irq_init` function from the init/main.c.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- General Protection Fault
- opcode
- Non-Maskable
- BOUND instruction
- CPU socket
- Interrupt Descriptor Table
- Interrupt Stack Table
- Paravirtualization
- .rept
- SIMD
- Coprocessor
- x86_64
- iret
- page fault
- breakpoint
- Global Descriptor Table
- stack frame
- Model Specific regiser
- percpu
- RCU
- MPX
- x87 FPU
- Previous part

Interrupts and Interrupt Handling. Part 7.

Introduction to external interrupts

This is the seventh part of the Interrupts and Interrupt Handling in the Linux kernel chapter and in the previous part we have finished with the exceptions which are generated by the processor. In this part we will continue to dive to the interrupt handling and will start with the external hardware interrupt handling. As you can remember, in the previous part we have finished with the `trap_init` function from the arch/x86/kernel/trap.c and the next step is the call of the `early_irq_init` function from the init/main.c.

Interrupts are signal that are sent across IRQ or `Interrupt Request Line` by a hardware or software. External hardware interrupts allow devices like keyboard, mouse and etc, to indicate that it needs attention of the processor. Once the processor receives the `Interrupt Request`, it will temporary stop execution of the running program and invoke special routine which depends on an interrupt. We already know that this routine is called interrupt handler (or how we will call it `ISR` or `Interrupt Service Routine` from this part). The `ISR` or `Interrupt Handler Routine` can be found in Interrupt Vector table that is located at fixed address in the memory. After the interrupt is handled processor resumes the interrupted process. At the boot/initialization time, the Linux kernel identifies all devices in the machine, and appropriate interrupt handlers are loaded into the interrupt table. As we saw in the previous parts, most exceptions are handled simply by the sending a Unix signal to the interrupted process. That's why kernel is can handle an exception quickly. Unfortunately we can not use this approach for the external hardware interrupts, because often they arrive after (and sometimes long after) the process to which they are related has been suspended. So it would make no sense to send a Unix signal to the current process. External interrupt handling depends on the type of an interrupt:

- `I/O` interrupts;
- Timer interrupts;
- Interprocessor interrupts.

I will try to describe all types of interrupts in this book.

Generally, a handler of an `I/O` interrupt must be flexible enough to service several devices at the same time. For example in the PCI bus architecture several devices may share the same `IRQ` line. In the simplest way the Linux kernel must do following thing when an `I/O` interrupt occurred:

- Save the value of an `IRQ` and the register's contents on the kernel stack;
- Send an acknowledgment to the hardware controller which is servicing the `IRQ` line;
- Execute the interrupt service routine (next we will call it `ISR`) which is associated with the device;
- Restore registers and return from an interrupt;

Ok, we know a little theory and now let's start with the `early_irq_init` function. The implementation of the `early_irq_init` function is in the kernel/irq/irqdesc.c. This function make early initialization of the `irq_desc` structure. The `irq_desc` structure is the foundation of interrupt management code in the Linux kernel. An array of this structure, which has the same name - `irq_desc`, keeps track of every interrupt request source in the Linux kernel. This structure defined in the include/linux/irqdesc.h and as you can note it depends on the `CONFIG_SPARSE_IRQ` kernel configuration option. This kernel configuration option enables support for sparse irqs. The `irq_desc` structure contains many different files:

- `irq_common_data` - per irq and chip data passed down to chip functions;
- `status_use_accessors` - contains status of the interrupt source which is combination of the values from the `enum` from the include/linux/irq.h and different macros which are defined in the same source code file;
- `kstat_irqs` - irq stats per-cpu;
- `handle_irq` - highlevel irq-events handler;
- `action` - identifies the interrupt service routines to be invoked when the IRQ occurs;

* `irq_count` - counter of interrupt occurrences on the IRQ line;
* `depth` - `0` if the IRQ line is enabled and a positive value if it has been disabled at least once;
* `last_unhandled` - aging timer for unhandled count;
* `irqs_unhandled` - count of the unhandled interrupts;
* `lock` - a spin lock used to serialize the accesses to the `IRQ` descriptor;
* `pending_mask` - pending rebalanced interrupts;
* `owner` - an owner of interrupt descriptor. Interrupt descriptors can be allocated from modules. This field is need to proved refcount on the module which provides the interrupts;
* and etc.

Of course it is not all fields of the `irq_desc` structure, because it is too long to describe each field of this structure, but we will see it all soon. Now let's start to dive into the implementation of the `early_irq_init` function.

Early external interrupts initialization

Now, let's look on the implementation of the `early_irq_init` function. Note that implementation of the `early_irq_init` function depends on the `CONFIG_SPARSE_IRQ` kernel configuration option. Now we consider implementation of the `early_irq_init` function when the `CONFIG_SPARSE_IRQ` kernel configuration option is not set. This function starts from the declaration of the following variables: `irq` descriptors counter, loop counter, memory node and the `irq_desc` descriptor:

```
int __init early_irq_init(void)
{
        int count, i, node = first_online_node;
        struct irq_desc *desc;
        ...
        ...
        ...
}
```

The `node` is an online NUMA node which depends on the `MAX_NUMNODES` value which depends on the `CONFIG_NODES_SHIFT` kernel configuration parameter:

```
#define MAX_NUMNODES    (1 << NODES_SHIFT)
...
...
...
#ifdef CONFIG_NODES_SHIFT
    #define NODES_SHIFT     CONFIG_NODES_SHIFT
#else
    #define NODES_SHIFT     0
#endif
```

As I already wrote, implementation of the `first_online_node` macro depends on the `MAX_NUMNODES` value:

```
#if MAX_NUMNODES > 1
  #define first_online_node       first_node(node_states[N_ONLINE])
#else
  #define first_online_node       0
```

The `node_states` is the enum which defined in the include/linux/nodemask.h and represent the set of the states of a node. In our case we are searching an online node and it will be `0` if `MAX_NUMNODES` is one or zero. If the `MAX_NUMNODES` is greater than one, the `node_states[N_ONLINE]` will return `1` and the `first_node` macro will be expands to the call of the `__first_node` function which will return `minimal` or the first online node:

```
#define first_node(src) __first_node(&(src))
```

```
static inline int __first_node(const nodemask_t *srcp)
{
        return min_t(int, MAX_NUMNODES, find_first_bit(srcp->bits, MAX_NUMNODES));
}
```

More about this will be in the another chapter about the `NUMA` . The next step after the declaration of these local variables is the call of the:

```
init_irq_default_affinity();
```

function. The `init_irq_default_affinity` function defined in the same source code file and depends on the `CONFIG_SMP` kernel configuration option allocates a given cpumask structure (in our case it is the `irq_default_affinity`):

```
#if defined(CONFIG_SMP)
cpumask_var_t irq_default_affinity;

static void __init init_irq_default_affinity(void)
{
        alloc_cpumask_var(&irq_default_affinity, GFP_NOWAIT);
        cpumask_setall(irq_default_affinity);
}
#else
static void __init init_irq_default_affinity(void)
{
}
#endif
```

We know that when a hardware, such as disk controller or keyboard, needs attention from the processor, it throws an interrupt. The interrupt tells to the processor that something has happened and that the processor should interrupt current process and handle an incoming event. In order to prevent multiple devices from sending the same interrupts, the IRQ system was established where each device in a computer system is assigned its own special IRQ so that its interrupts are unique. Linux kernel can assign certain `IRQs` to specific processors. This is known as `SMP IRQ affinity` , and it allows you control how your system will respond to various hardware events (that's why it has certain implementation only if the `CONFIG_SMP` kernel configuration option is set). After we allocated `irq_default_affinity` cpumask, we can see `printk` output:

```
printk(KERN_INFO "NR_IRQS:%d\n", NR_IRQS);
```

which prints `NR_IRQS` :

```
~$ dmesg | grep NR_IRQS
[    0.000000] NR_IRQS:4352
```

The `NR_IRQS` is the maximum number of the `irq` descriptors or in another words maximum number of interrupts. Its value depends on the state of the `CONFIG_X86_IO_APIC` kernel configuration option. If the `CONFIG_X86_IO_APIC` is not set and the Linux kernel uses an old PIC chip, the `NR_IRQS` is:

```
#define NR_IRQS_LEGACY                    16

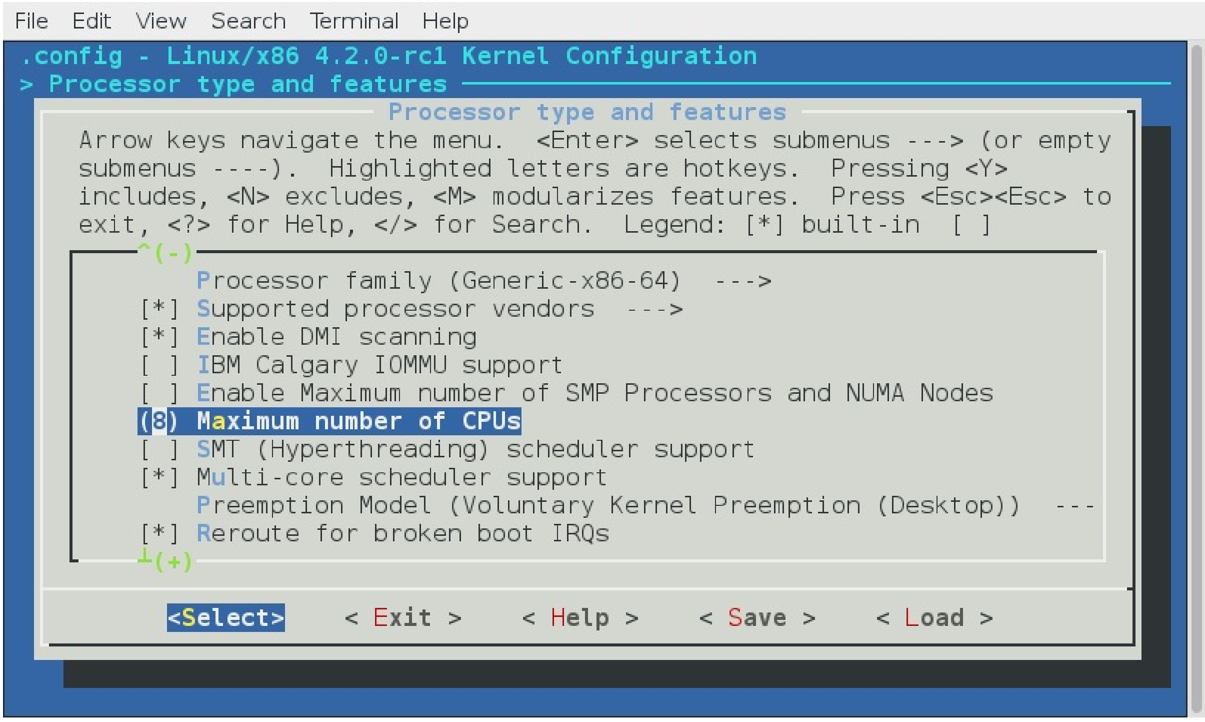
#ifdef CONFIG_X86_IO_APIC
...
...
...
#else
# define NR_IRQS                        NR_IRQS_LEGACY
#endif
```

In other way, when the `CONFIG_X86_IO_APIC` kernel configuration option is set, the `NR_IRQS` depends on the amount of the processors and amount of the interrupt vectors:

```
#define CPU_VECTOR_LIMIT               (64 * NR_CPUS)
#define NR_VECTORS                     256
#define IO_APIC_VECTOR_LIMIT           ( 32 * MAX_IO_APICS )
#define MAX_IO_APICS                   128

# define NR_IRQS                                        \
        (CPU_VECTOR_LIMIT > IO_APIC_VECTOR_LIMIT ?      \
                (NR_VECTORS + CPU_VECTOR_LIMIT)  :      \
                (NR_VECTORS + IO_APIC_VECTOR_LIMIT))
...
...
...
```

We remember from the previous parts, that the amount of processors we can set during Linux kernel configuration process with the `CONFIG_NR_CPUS` configuration option:



In the first case (`CPU_VECTOR_LIMIT > IO_APIC_VECTOR_LIMIT`), the `NR_IRQS` will be `4352` , in the second case (`CPU_VECTOR_LIMIT < IO_APIC_VECTOR_LIMIT`), the `NR_IRQS` will be `768` . In my case the `NR_CPUS` is `8` as you can see in the my configuration, the `CPU_VECTOR_LIMIT` is `512` and the `IO_APIC_VECTOR_LIMIT` is `4096` . So `NR_IRQS` for my configuration is `4352` :

```
~$ dmesg | grep NR_IRQS
[    0.000000] NR_IRQS:4352
```

In the next step we assign array of the IRQ descriptors to the `irq_desc` variable which we defined in the start of the `early_irq_init` function and calculate count of the `irq_desc` array with the `ARRAY_SIZE` macro:

```
desc = irq_desc;
count = ARRAY_SIZE(irq_desc);
```

The `irq_desc` array defined in the same source code file and looks like:

```
struct irq_desc irq_desc[NR_IRQS] __cacheline_aligned_in_smp = {
        [0 ... NR_IRQS-1] = {
                .handle_irq     = handle_bad_irq,
                .depth          = 1,
                .lock           = __RAW_SPIN_LOCK_UNLOCKED(irq_desc->lock),
        }
};
```

The `irq_desc` is array of the `irq` descriptors. It has three already initialized fields:

- `handle_irq` - as I already wrote above, this field is the highlevel irq-event handler. In our case it initialized with the `handle_bad_irq` function that defined in the kernel/irq/handle.c source code file and handles spurious and unhandled irqs;
- `depth` - `0` if the IRQ line is enabled and a positive value if it has been disabled at least once;
- `lock` - A spin lock used to serialize the accesses to the `IRQ` descriptor.

As we calculated count of the interrupts and initialized our `irq_desc` array, we start to fill descriptors in the loop:

```
for (i = 0; i < count; i++) {
    desc[i].kstat_irqs = alloc_percpu(unsigned int);
    alloc_masks(&desc[i], GFP_KERNEL, node);
    raw_spin_lock_init(&desc[i].lock);
    lockdep_set_class(&desc[i].lock, &irq_desc_lock_class);
    desc_set_defaults(i, &desc[i], node, NULL);
}
```

We are going through the all interrupt descriptors and do the following things:

First of all we allocate percpu variable for the `irq` kernel statistic with the `alloc_percpu` macro. This macro allocates one instance of an object of the given type for every processor on the system. You can access kernel statistic from the userspace via `/proc/stat` :

```
~$ cat /proc/stat
cpu  207907 68 53904 5427850 14394 0 394 0 0 0
cpu0 25881 11 6684 679131 1351 0 18 0 0 0
cpu1 24791 16 5894 679994 2285 0 24 0 0 0
cpu2 26321 4 7154 678924 664 0 71 0 0 0
cpu3 26648 8 6931 678891 414 0 244 0 0 0
...
...
...
```

Where the sixth column is the servicing interrupts. After this we allocate cpumask for the given irq descriptor affinity and initialize the spinlock for the given interrupt descriptor. After this before the critical section, the lock will be acquired with a call of the `raw_spin_lock` and unlocked with the call of the `raw_spin_unlock` . In the next step we call the `lockdep_set_class` macro which set the Lock validator `irq_desc_lock_class` class for the lock of the given interrupt descriptor. More about `lockdep` , `spinlock` and other synchronization primitives will be described in the separate chapter.

In the end of the loop we call the `desc_set_defaults` function from the kernel/irq/irqdesc.c. This function takes four parameters:

- number of a irq;
- interrupt descriptor;
- online `NUMA` node;
- owner of interrupt descriptor. Interrupt descriptors can be allocated from modules. This field is need to proved refcount on the module which provides the interrupts;

and fills the rest of the `irq_desc` fields. The `desc_set_defaults` function fills interrupt number, `irq` chip, platform-specific per-chip private data for the chip methods, per-IRQ data for the `irq_chip` methods and MSI descriptor for the per `irq` and `irq` chip data:

```
desc->irq_data.irq = irq;
desc->irq_data.chip = &no_irq_chip;
desc->irq_data.chip_data = NULL;
desc->irq_data.handler_data = NULL;
desc->irq_data.msi_desc = NULL;
...
...
...
```

The `irq_data.chip` structure provides general `API` like the `irq_set_chip`, `irq_set_irq_type` and etc, for the irq controller drivers. You can find it in the kernel/irq/chip.c source code file.

After this we set the status of the accessor for the given descriptor and set disabled state of the interrupts:

```
...
...
...
irq_settings_clr_and_set(desc, ~0, _IRQ_DEFAULT_INIT_FLAGS);
irqd_set(&desc->irq_data, IRQD_IRQ_DISABLED);
...
...
...
```

In the next step we set the high level interrupt handlers to the `handle_bad_irq` which handles spurious and unhandled irqs (as the hardware stuff is not initialized yet, we set this handler), set `irq_desc.desc` to `1` which means that an `IRQ` is disabled, reset count of the unhandled interrupts and interrupts in general:

```
...
...
...
desc->handle_irq = handle_bad_irq;
desc->depth = 1;
desc->irq_count = 0;
desc->irqs_unhandled = 0;
desc->name = NULL;
desc->owner = owner;
...
...
...
```

After this we go through the all possible processor with the for_each_possible_cpu helper and set the `kstat_irqs` to zero for the given interrupt descriptor:

```
    for_each_possible_cpu(cpu)
        *per_cpu_ptr(desc->kstat_irqs, cpu) = 0;
```

and call the `desc_smp_init` function from the kernel/irq/irqdesc.c that initializes `NUMA` node of the given interrupt descriptor, sets default `SMP` affinity and clears the `pending_mask` of the given interrupt descriptor depends on the value of the `CONFIG_GENERIC_PENDING_IRQ` kernel configuration option:

```
static void desc_smp_init(struct irq_desc *desc, int node)
{
        desc->irq_data.node = node;
        cpumask_copy(desc->irq_data.affinity, irq_default_affinity);
#ifdef CONFIG_GENERIC_PENDING_IRQ
        cpumask_clear(desc->pending_mask);
#endif
}
```

In the end of the `early_irq_init` function we return the return value of the `arch_early_irq_init` function:

```
return arch_early_irq_init();
```

This function defined in the kernel/apic/vector.c and contains only one call of the `arch_early_ioapic_init` function from the kernel/apic/io_apic.c. As we can understand from the `arch_early_ioapic_init` function's name, this function makes early initialization of the I/O APIC. First of all it make a check of the number of the legacy interrupts with the call of the `nr_legacy_irqs` function. If we have no legacy interrupts with the Intel 8259 programmable interrupt controller we set `io_apic_irqs` to the `0xffffffffffffffff` :

```
if (!nr_legacy_irqs())
    io_apic_irqs = ~0UL;
```

After this we are going through the all `I/O APICs` and allocate space for the registers with the call of the `alloc_ioapic_saved_registers` :

```
for_each_ioapic(i)
    alloc_ioapic_saved_registers(i);
```

And in the end of the `arch_early_ioapic_init` function we are going through the all legacy irqs (from `IRQ0` to `IRQ15`) in the loop and allocate space for the `irq_cfg` which represents configuration of an irq on the given `NUMA` node:

```
for (i = 0; i < nr_legacy_irqs(); i++) {
    cfg = alloc_irq_and_cfg_at(i, node);
    cfg->vector = IRQ0_VECTOR + i;
    cpumask_setall(cfg->domain);
}
```

That's all.

Sparse IRQs

We already saw in the beginning of this part that implementation of the `early_irq_init` function depends on the `CONFIG_SPARSE_IRQ` kernel configuration option. Previously we saw implementation of the `early_irq_init` function when the `CONFIG_SPARSE_IRQ` configuration option is not set, now let's look on the its implementation when this option is set. Implementation of this function very similar, but little differ. We can see the same definition of variables and call of the `init_irq_default_affinity` in the beginning of the `early_irq_init` function:

```
#ifdef CONFIG_SPARSE_IRQ
int __init early_irq_init(void)
{
    int i, initcnt, node = first_online_node;
    struct irq_desc *desc;

    init_irq_default_affinity();
    ...
    ...
    ...
}
#else
...
...
...
```

But after this we can see the following call:

```
initcnt = arch_probe_nr_irqs();
```

The `arch_probe_nr_irqs` function defined in the arch/x86/kernel/apic/vector.c and calculates count of the pre-allocated irqs and update `nr_irqs` with its number. But stop. Why there are pre-allocated irqs? There is alternative form of interrupts called - Message Signaled Interrupts available in the PCI. Instead of assigning a fixed number of the interrupt request, the device is allowed to record a message at a particular address of RAM, in fact, the display on the Local APIC. `MSI` permits a device to allocate `1`, `2`, `4`, `8`, `16` or `32` interrupts and `MSI-X` permits a device to allocate up to `2048` interrupts. Now we know that irqs can be pre-allocated. More about `MSI` will be in a next part, but now let's look on the `arch_probe_nr_irqs` function. We can see the check which assign amount of the interrupt vectors for the each processor in the system to the `nr_irqs` if it is greater and calculate the `nr` which represents number of `MSI` interrupts:

```
int nr_irqs = NR_IRQS;

if (nr_irqs > (NR_VECTORS * nr_cpu_ids))
    nr_irqs = NR_VECTORS * nr_cpu_ids;

nr = (gsi_top + nr_legacy_irqs()) + 8 * nr_cpu_ids;
```

Take a look on the `gsi_top` variable. Each `APIC` is identified with its own `ID` and with the offset where its `IRQ` starts. It is called `GSI` base or `Global System Interrupt` base. So the `gsi_top` represents it. We get the `Global System Interrupt` base from the MultiProcessor Configuration Table table (you can remember that we have parsed this table in the sixth part of the Linux Kernel initialization process chapter).

After this we update the `nr` depends on the value of the `gsi_top`:

```
#if defined(CONFIG_PCI_MSI) || defined(CONFIG_HT_IRQ)
        if (gsi_top <= NR_IRQS_LEGACY)
                nr +=  8 * nr_cpu_ids;
        else
                nr += gsi_top * 16;
#endif
```

Update the `nr_irqs` if it less than `nr` and return the number of the legacy irqs:

```
if (nr < nr_irqs)
    nr_irqs = nr;

return nr_legacy_irqs();
}
```

The next after the `arch_probe_nr_irqs` is printing information about number of `IRQs`:

```
printk(KERN_INFO "NR_IRQS:%d nr_irqs:%d %d\n", NR_IRQS, nr_irqs, initcnt);
```

We can find it in the dmesg output:

```
$ dmesg | grep NR_IRQS
[    0.000000] NR_IRQS:4352 nr_irqs:488 16
```

After this we do some checks that `nr_irqs` and `initcnt` values is not greater than maximum allowable number of `irqs`:

```
if (WARN_ON(nr_irqs > IRQ_BITMAP_BITS))
    nr_irqs = IRQ_BITMAP_BITS;

if (WARN_ON(initcnt > IRQ_BITMAP_BITS))
```

```
    initcnt = IRQ_BITMAP_BITS;
```

where `IRQ_BITMAP_BITS` is equal to the `NR_IRQS` if the `CONFIG_SPARSE_IRQ` is not set and `NR_IRQS + 8196` in other way. In the next step we are going over all interrupt descriptors which need to be allocated in the loop and allocate space for the descriptor and insert to the `irq_desc_tree` radix tree:

```
for (i = 0; i < initcnt; i++) {
    desc = alloc_desc(i, node, NULL);
    set_bit(i, allocated_irqs);
    irq_insert_desc(i, desc);
}
```

In the end of the `early_irq_init` function we return the value of the call of the `arch_early_irq_init` function as we did it already in the previous variant when the `CONFIG_SPARSE_IRQ` option was not set:

```
return arch_early_irq_init();
```

That's all.

Conclusion

It is the end of the seventh part of the Interrupts and Interrupt Handling chapter and we started to dive into external hardware interrupts in this part. We saw early initialization of the `irq_desc` structure which represents description of an external interrupt and contains information about it like list of irq actions, information about interrupt handler, interrupt's owner, count of the unhandled interrupt and etc. In the next part we will continue to research external interrupts.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- IRQ
- numa
- Enum type
- cpumask
- percpu
- spinlock
- critical section
- Lock validator
- MSI
- I/O APIC
- Local APIC
- Intel 8259
- PIC
- MultiProcessor Configuration Table
- radix tree
- dmesg

Interrupts and Interrupt Handling. Part 8.

Non-early initialization of the IRQs

This is the eighth part of the Interrupts and Interrupt Handling in the Linux kernel chapter and in the previous part we started to dive into the external hardware interrupts. We looked on the implementation of the `early_irq_init` function from the kernel/irq/irqdesc.c source code file and saw the initialization of the `irq_desc` structure in this function. Remind that `irq_desc` structure (defined in the include/linux/irqdesc.h is the foundation of interrupt management code in the Linux kernel and represents an interrupt descriptor. In this part we will continue to dive into the initialization stuff which is related to the external hardware interrupts.

Right after the call of the `early_irq_init` function in the init/main.c we can see the call of the `init_IRQ` function. This function is architecture-specific and defined in the arch/x86/kernel/irqinit.c. The `init_IRQ` function makes initialization of the `vector_irq` percpu variable that defined in the same arch/x86/kernel/irqinit.c source code file:

```
...
DEFINE_PER_CPU(vector_irq_t, vector_irq) = {
         [0 ... NR_VECTORS - 1] = -1,
};
...
```

and represents `percpu` array of the interrupt vector numbers. The `vector_irq_t` defined in the arch/x86/include/asm/hw_irq.h and expands to the:

```
typedef int vector_irq_t[NR_VECTORS];
```

where `NR_VECTORS` is count of the vector number and as you can remember from the first part of this chapter it is `256` for the x86_64:

```
#define NR_VECTORS                       256
```

So, in the start of the `init_IRQ` function we fill the `vector_irq` percpu array with the vector number of the `legacy` interrupts:

```
void __init init_IRQ(void)
{
    int i;

    for (i = 0; i < nr_legacy_irqs(); i++)
        per_cpu(vector_irq, 0)[IRQ0_VECTOR + i] = i;
...
...
...
}
```

This `vector_irq` will be used during the first steps of an external hardware interrupt handling in the `do_IRQ` function from the arch/x86/kernel/irq.c:

```
__visible unsigned int __irq_entry do_IRQ(struct pt_regs *regs)
{
    ...
    ...
    ...
    irq = __this_cpu_read(vector_irq[vector]);
```

```
    if (!handle_irq(irq, regs)) {
        ...
        ...
        ...
    }

    exiting_irq();
    ...
    ...
    return 1;
}
```

Why is `legacy` here? Actually all interrupts are handled by the modern IO-APIC controller. But these interrupts (from `0x30` to `0x3f`) by legacy interrupt-controllers like Programmable Interrupt Controller. If these interrupts are handled by the `I/O APIC` then this vector space will be freed and re-used. Let's look on this code closer. First of all the `nr_legacy_irqs` defined in the arch/x86/include/asm/i8259.h and just returns the `nr_legacy_irqs` field from the `legacy_pic` structure:

```
static inline int nr_legacy_irqs(void)
{
        return legacy_pic->nr_legacy_irqs;
}
```

This structure defined in the same header file and represents non-modern programmable interrupts controller:

```
struct legacy_pic {
        int nr_legacy_irqs;
        struct irq_chip *chip;
        void (*mask)(unsigned int irq);
        void (*unmask)(unsigned int irq);
        void (*mask_all)(void);
        void (*restore_mask)(void);
        void (*init)(int auto_eoi);
        int (*irq_pending)(unsigned int irq);
        void (*make_irq)(unsigned int irq);
};
```

Actual default maximum number of the legacy interrupts represented by the `NR_IRQ_LEGACY` macro from the arch/x86/include/asm/irq_vectors.h:

```
#define NR_IRQS_LEGACY                    16
```

In the loop we are accessing the `vecto_irq` per-cpu array with the `per_cpu` macro by the `IRQ0_VECTOR + i` index and write the legacy vector number there. The `IRQ0_VECTOR` macro defined in the arch/x86/include/asm/irq_vectors.h header file and expands to the `0x30`:

```
#define FIRST_EXTERNAL_VECTOR           0x20

#define IRQ0_VECTOR                     ((FIRST_EXTERNAL_VECTOR + 16) & ~15)
```

Why is `0x30` here? You can remember from the first part of this chapter that first 32 vector numbers from `0` to `31` are reserved by the processor and used for the processing of architecture-defined exceptions and interrupts. Vector numbers from `0x30` to `0x3f` are reserved for the ISA. So, it means that we fill the `vector_irq` from the `IRQ0_VECTOR` which is equal to the `32` to the `IRQ0_VECTOR + 16` (before the `0x30`).

In the end of the `init_IRQ` function we can see the call of the following function:

```
x86_init.irqs.intr_init();
```

from the arch/x86/kernel/x86_init.c source code file. If you have read chapter about the Linux kernel initialization process, you can remember the `x86_init` structure. This structure contains a couple of files which are points to the function related to the platform setup (`x86_64` in our case), for example `resources` - related with the memory resources, `mpparse` - related with the parsing of the MultiProcessor Configuration Table table and etc.). As we can see the `x86_init` also contains the `irqs` field which contains three following fields:

```
struct x86_init_ops x86_init __initdata
{
    ...
    ...
    ...
    .irqs = {
                .pre_vector_init        = init_ISA_irqs,
                .intr_init              = native_init_IRQ,
                .trap_init              = x86_init_noop,
    },
    ...
    ...
    ...
}
```

Now, we are interesting in the `native_init_IRQ` . As we can note, the name of the `native_init_IRQ` function contains the `native_` prefix which means that this function is architecture-specific. It defined in the arch/x86/kernel/irqinit.c and executes general initialization of the Local APIC and initialization of the ISA irqs. Let's look on the implementation of the `native_init_IRQ` function and will try to understand what occurs there. The `native_init_IRQ` function starts from the execution of the following function:

```
x86_init.irqs.pre_vector_init();
```

As we can see above, the `pre_vector_init` points to the `init_ISA_irqs` function that defined in the same source code file and as we can understand from the function's name, it makes initialization of the `ISA` related interrupts. The `init_ISA_irqs` function starts from the definition of the `chip` variable which has a `irq_chip` type:

```
void __init init_ISA_irqs(void)
{
    struct irq_chip *chip = legacy_pic->chip;
    ...
    ...
    ...
```

The `irq_chip` structure defined in the include/linux/irq.h header file and represents hardware interrupt chip descriptor. It contains:

* `name` - name of a device. Used in the `/proc/interrupts` :

```
$ cat /proc/interrupts
           CPU0       CPU1       CPU2       CPU3       CPU4       CPU5       CPU6       CPU7
  0:         16          0          0          0          0          0          0          0   IO-APIC   2-edge
    timer
  1:          2          0          0          0          0          0          0          0   IO-APIC   1-edge
    i8042
  8:          1          0          0          0          0          0          0          0   IO-APIC   8-edge
    rtc0
```

look on the last column;

* `(*irq_mask)(struct irq_data *data)` - mask an interrupt source;

- `(*irq_ack)(struct irq_data *data)` - start of a new interrupt;
- `(*irq_startup)(struct irq_data *data)` - start up the interrupt;
- `(*irq_shutdown)(struct irq_data *data)` - shutdown the interrupt
- and etc.

fields. Note that the `irq_data` structure represents set of the per irq chip data passed down to chip functions. It contains `mask` - precomputed bitmask for accessing the chip registers, `irq` - interrupt number, `hwirq` - hardware interrupt number, local to the interrupt domain chip low level interrupt hardware access and etc.

After this depends on the `CONFIG_X86_64` and `CONFIG_X86_LOCAL_APIC` kernel configuration option call the `init_bsp_APIC` function from the arch/x86/kernel/apic/apic.c:

```
#if defined(CONFIG_X86_64) || defined(CONFIG_X86_LOCAL_APIC)
    init_bsp_APIC();
#endif
```

This function makes initialization of the APIC of `bootstrap processor` (or processor which starts first). It starts from the check that we found SMP config (read more about it in the sixth part of the Linux kernel initialization process chapter) and the processor has `APIC` :

```
if (smp_found_config || !cpu_has_apic)
    return;
```

In other way we return from this function. In the next step we call the `clear_local_APIC` function from the same source code file that shutdowns the local `APIC` (more about it will be in the chapter about the `Advanced Programmable Interrupt Controller`) and enable `APIC` of the first processor by the setting `unsigned int value` to the `APIC_SPIV_APIC_ENABLED` :

```
value = apic_read(APIC_SPIV);
value &= ~APIC_VECTOR_MASK;
value |= APIC_SPIV_APIC_ENABLED;
```

and writing it with the help of the `apic_write` function:

```
apic_write(APIC_SPIV, value);
```

After we have enabled `APIC` for the bootstrap processor, we return to the `init_ISA_irqs` function and in the next step we initialize legacy `Programmable Interrupt Controller` and set the legacy chip and handler for the each legacy irq:

```
legacy_pic->init(0);

for (i = 0; i < nr_legacy_irqs(); i++)
    irq_set_chip_and_handler(i, chip, handle_level_irq);
```

Where can we find `init` function? The `legacy_pic` defined in the arch/x86/kernel/i8259.c and it is:

```
struct legacy_pic *legacy_pic = &default_legacy_pic;
```

Where the `default_legacy_pic` is:

```
struct legacy_pic default_legacy_pic = {
    ...
    ...
    ...
    .init = init_8259A,
    ...
```

```
    ...
    ...
}
```

The `init_8259A` function defined in the same source code file and executes initialization of the Intel 8259 `` `Programmable Interrupt Controller `` (more about it will be in the separate chapter about `Programmable Interrupt Controllers` and `APIC`).

Now we can return to the `native_init_IRQ` function, after the `init_ISA_irqs` function finished its work. The next step is the call of the `apic_intr_init` function that allocates special interrupt gates which are used by the SMP architecture for the Inter-processor interrupt. The `alloc_intr_gate` macro from the arch/x86/include/asm/desc.h used for the interrupt descriptor allocation:

```
#define alloc_intr_gate(n, addr)                        \
do {                                                    \
        alloc_system_vector(n);                         \
        set_intr_gate(n, addr);                         \
} while (0)
```

As we can see, first of all it expands to the call of the `alloc_system_vector` function that checks the given vector number in the `used_vectors` bitmap (read previous part about it) and if it is not set in the `used_vectors` bitmap we set it. After this we test that the `first_system_vector` is greater than given interrupt vector number and if it is greater we assign it:

```
if (!test_bit(vector, used_vectors)) {
    set_bit(vector, used_vectors);
    if (first_system_vector > vector)
        first_system_vector = vector;
} else {
    BUG();
}
```

We already saw the `set_bit` macro, now let's look on the `test_bit` and the `first_system_vector` . The first `test_bit` macro defined in the arch/x86/include/asm/bitops.h and looks like this:

```
#define test_bit(nr, addr)                      \
        (__builtin_constant_p((nr))             \
         ? constant_test_bit((nr), (addr))      \
         : variable_test_bit((nr), (addr)))
```

We can see the ternary operator here make a test with the gcc built-in function `__builtin_constant_p` tests that given vector number (`nr`) is known at compile time. If you're feeling misunderstanding of the `__builtin_constant_p` , we can make simple test:

```
#include <stdio.h>

#define PREDEFINED_VAL 1

int main() {
    int i = 5;
    printf("__builtin_constant_p(i) is %d\n", __builtin_constant_p(i));
    printf("__builtin_constant_p(PREDEFINED_VAL) is %d\n", __builtin_constant_p(PREDEFINED_VAL));
    printf("__builtin_constant_p(100) is %d\n", __builtin_constant_p(100));

    return 0;
}
```

and look on the result:

```
$ gcc test.c -o test
```

```
$ ./test
__builtin_constant_p(i) is 0
__builtin_constant_p(PREDEFINED_VAL) is 1
__builtin_constant_p(100) is 1
```

Now I think it must be clear for you. Let's get back to the `test_bit` macro. If the `__builtin_constant_p` will return non-zero, we call `constant_test_bit` function:

```
static inline int constant_test_bit(int nr, const void *addr)
{
    const u32 *p = (const u32 *)addr;

    return ((1UL << (nr & 31)) & (p[nr >> 5])) != 0;
}
```

and the `variable_test_bit` in other way:

```
static inline int variable_test_bit(int nr, const void *addr)
{
        u8 v;
        const u32 *p = (const u32 *)addr;

        asm("btl %2,%1; setc %0" : "=qm" (v) : "m" (*p), "Ir" (nr));
        return v;
}
```

What's the difference between two these functions and why do we need in two different functions for the same purpose? As you already can guess main purpose is optimization. If we will write simple example with these functions:

```
#define CONST 25

int main() {
    int nr = 24;
    variable_test_bit(nr, (int*)0x10000000);
    constant_test_bit(CONST, (int*)0x10000000)
    return 0;
}
```

and will look on the assembly output of our example we will see following assembly code:

```
pushq    %rbp
movq    %rsp, %rbp

movl    $268435456, %esi
movl    $25, %edi
call    constant_test_bit
```

for the `constant_test_bit` , and:

```
pushq    %rbp
movq    %rsp, %rbp

subq    $16, %rsp
movl    $24, -4(%rbp)
movl    -4(%rbp), %eax
movl    $268435456, %esi
movl    %eax, %edi
call    variable_test_bit
```

for the `variable_test_bit` . These two code listings starts with the same part, first of all we save base of the current stack frame in the `%rbp` register. But after this code for both examples is different. In the first example we put `$268435456` (here the `$268435456` is our second parameter - `0x10000000`) to the `esi` and `$25` (our first parameter) to the `edi` register and call `constant_test_bit` . We put function parameters to the `esi` and `edi` registers because as we are learning Linux kernel for the `x86_64` architecture we use `System V AMD64 ABI` calling convention. All is pretty simple. When we are using predefined constant, the compiler can just substitute its value. Now let's look on the second part. As you can see here, the compiler can not substitute value from the `nr` variable. In this case compiler must calculate its offset on the program's stack frame. We subtract `16` from the `rsp` register to allocate stack for the local variables data and put the `$24` (value of the `nr` variable) to the `rbp` with offset `-4` . Our stack frame will be like this:

```
         <- stack grows

             %[rbp]
                |
+----------+ +---------+ +---------+ +--------+
|          | |         | | return  | |        |
|    nr    |-|         |-|         |-|  argc  |
|          | |         | | address | |        |
+----------+ +---------+ +---------+ +--------+
                |
             %[rsp]
```

After this we put this value to the `eax` , so `eax` register now contains value of the `nr` . In the end we do the same that in the first example, we put the `$268435456` (the first parameter of the `variable_test_bit` function) and the value of the `eax` (value of `nr`) to the `edi` register (the second parameter of the `variable_test_bit function`).

The next step after the `apic_intr_init` function will finish its work is the setting interrupt gates from the `FIRST_EXTERNAL_VECTOR` or `0x20` to the `0x256` :

```
i = FIRST_EXTERNAL_VECTOR;

#ifndef CONFIG_X86_LOCAL_APIC
#define first_system_vector NR_VECTORS
#endif

for_each_clear_bit_from(i, used_vectors, first_system_vector) {
    set_intr_gate(i, irq_entries_start + 8 * (i - FIRST_EXTERNAL_VECTOR));
}
```

But as we are using the `for_each_clear_bit_from` helper, we set only non-initialized interrupt gates. After this we use the same `for_each_clear_bit_from` helper to fill the non-filled interrupt gates in the interrupt table with the `spurious_interrupt` :

```
#ifdef CONFIG_X86_LOCAL_APIC
for_each_clear_bit_from(i, used_vectors, NR_VECTORS)
    set_intr_gate(i, spurious_interrupt);
#endif
```

Where the `spurious_interrupt` function represent interrupt handler for the `spurious` interrupt. Here the `used_vectors` is the `unsigned long` that contains already initialized interrupt gates. We already filled first `32` interrupt vectors in the `trap_init` function from the arch/x86/kernel/setup.c source code file:

```
for (i = 0; i < FIRST_EXTERNAL_VECTOR; i++)
    set_bit(i, used_vectors);
```

You can remember how we did it in the sixth part of this chapter.

In the end of the `native_init_IRQ` function we can see the following check:

```
if (!acpi_ioapic && !of_ioapic && nr_legacy_irqs())
    setup_irq(2, &irq2);
```

First of all let's deal with the condition. The `acpi_ioapic` variable represents existence of I/O APIC. It defined in the arch/x86/kernel/acpi/boot.c. This variable set in the `acpi_set_irq_model_ioapic` function that called during the processing `Multiple APIC Description Table`. This occurs during initialization of the architecture-specific stuff in the arch/x86/kernel/setup.c (more about it we will know in the other chapter about APIC). Note that the value of the `acpi_ioapic` variable depends on the `CONFIG_ACPI` and `CONFIG_X86_LOCAL_APIC` Linux kernel configuration options. If these options did not set, this variable will be just zero:

```
#define acpi_ioapic 0
```

The second condition - `!of_ioapic && nr_legacy_irqs()` checks that we do not use Open Firmware `I/O APIC` and legacy interrupt controller. We already know about the `nr_legacy_irqs`. The second is `of_ioapic` variable defined in the arch/x86/kernel/devicetree.c and initialized in the `dtb_ioapic_setup` function that build information about `APICs` in the devicetree. Note that `of_ioapic` variable depends on the `CONFIG_OF` Linux kernel configuration option. If this option is not set, the value of the `of_ioapic` will be zero too:

```
#ifdef CONFIG_OF
extern int of_ioapic;
...
...
...
#else
#define of_ioapic 0
...
...
...
#endif
```

If the condition will return non-zero value we call the:

```
setup_irq(2, &irq2);
```

function. First of all about the `irq2`. The `irq2` is the `irqaction` structure that defined in the arch/x86/kernel/irqinit.c source code file and represents `IRQ 2` line that is used to query devices connected cascade:

```
static struct irqaction irq2 = {
    .handler = no_action,
    .name = "cascade",
    .flags = IRQF_NO_THREAD,
};
```

Some time ago interrupt controller consisted of two chips and one was connected to second. The second chip that was connected to the first chip via this `IRQ 2` line. This chip serviced lines from `8` to `15` and after this lines of the first chip. So, for example Intel 8259A has following lines:

- `IRQ 0` - system time;
- `IRQ 1` - keyboard;
- `IRQ 2` - used for devices which are cascade connected;
- `IRQ 8` - RTC;
- `IRQ 9` - reserved;
- `IRQ 10` - reserved;
- `IRQ 11` - reserved;

* `IRQ 12` - `ps/2` mouse;
* `IRQ 13` - coprocessor;
* `IRQ 14` - hard drive controller;
* `IRQ 1` - reserved;
* `IRQ 3` - `COM2` and `COM4` ;
* `IRQ 4` - `COM1` and `COM3` ;
* `IRQ 5` - `LPT2` ;
* `IRQ 6` - drive controller;
* `IRQ 7` - `LPT1` .

The `setup_irq` function defined in the kernel/irq/manage.c and takes two parameters:

* vector number of an interrupt;
* `irqaction` structure related with an interrupt.

This function initializes interrupt descriptor from the given vector number at the beginning:

```
struct irq_desc *desc = irq_to_desc(irq);
```

And call the `__setup_irq` function that setups given interrupt:

```
chip_bus_lock(desc);
retval = __setup_irq(irq, desc, act);
chip_bus_sync_unlock(desc);
return retval;
```

Note that the interrupt descriptor is locked during `__setup_irq` function will work. The `__setup_irq` function makes many different things: It creates a handler thread when a thread function is supplied and the interrupt does not nest into another interrupt thread, sets the flags of the chip, fills the `irqaction` structure and many many more.

All of the above it creates `/prov/vector_number` directory and fills it, but if you are using modern computer all values will be zero there:

```
$ cat /proc/irq/2/node
0

$cat /proc/irq/2/affinity_hint
00

cat /proc/irq/2/spurious
count 0
unhandled 0
last_unhandled 0 ms
```

because probably `APIC` handles interrupts on the our machine.

That's all.

Conclusion

It is the end of the eighth part of the Interrupts and Interrupt Handling chapter and we continued to dive into external hardware interrupts in this part. In the previous part we started to do it and saw early initialization of the `IRQs` . In this part we already saw non-early interrupts initialization in the `init_IRQ` function. We saw initialization of the `vector_irq` per-cpu array which is store vector numbers of the interrupts and will be used during interrupt handling and initialization of other stuff which is related to the external hardware interrupts.

In the next part we will continue to learn interrupts handling related stuff and will see initialization of the `softirqs`.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- IRQ
- percpu
- x86_64
- Intel 8259
- Programmable Interrupt Controller
- ISA
- MultiProcessor Configuration Table
- Local APIC
- I/O APIC
- SMP
- Inter-processor interrupt
- ternary operator
- gcc
- calling convention
- PDF. System V Application Binary Interface AMD64
- Call stack
- Open Firmware
- devicetree
- RTC
- Previous part

Interrupts and Interrupt Handling. Part 9.

Introduction to deferred interrupts (Softirq, Tasklets and Workqueues)

It is the nine part of the Interrupts and Interrupt Handling in the Linux kernel chapter and in the previous Previous part we saw implementation of the `init_IRQ` from that defined in the arch/x86/kernel/irqinit.c source code file. So, we will continue to dive into the initialization stuff which is related to the external hardware interrupts in this part.

Interrupts may have different important characteristics and there are two among them:

- Handler of an interrupt must execute quickly;
- Sometime an interrupt handler must do a large amount of work.

As you can understand, it is almost impossible to make so that both characteristics were valid. Because of these, previously the handling of interrupts was split into two parts:

- Top half;
- Bottom half;

Once the Linux kernel was one of the ways the organization postprocessing, and which was called: `the bottom half` of the processor, but now it is already not actual. Now this term has remained as a common noun referring to all the different ways of organizing deferred processing of an interrupt.The deferred processing of an interrupt suggests that some of the actions for an interrupt may be postponed to a later execution when the system will be less loaded. As you can suggests, an interrupt handler can do large amount of work that is impermissible as it executes in the context where interrupts are disabled. That's why processing of an interrupt can be split on two different parts. In the first part, the main handler of an interrupt does only minimal and the most important job. After this it schedules the second part and finishes its work. When the system is less busy and context of the processor allows to handle interrupts, the second part starts its work and finishes to process remaining part of a deferred interrupt.

There are three types of `deferred interrupts` in the Linux kernel:

- `softirqs` ;
- `tasklets` ;
- `workqueues` ;

And we will see description of all of these types in this part. As I said, we saw only a little bit about this theme, so, now is time to dive deep into details about this theme.

Softirqs

With the advent of parallelisms in the Linux kernel, all new schemes of implementation of the bottom half handlers are built on the performance of the processor specific kernel thread that called `ksoftirqd` (will be discussed below). Each processor has its own thread that is called `ksoftirqd/n` where the `n` is the number of the processor. We can see it in the output of the `systemd-cgls` util:

```
$ systemd-cgls -k | grep ksoft
├─   3 [ksoftirqd/0]
├─  13 [ksoftirqd/1]
├─  18 [ksoftirqd/2]
├─  23 [ksoftirqd/3]
├─  28 [ksoftirqd/4]
├─  33 [ksoftirqd/5]
├─  38 [ksoftirqd/6]
```

```
├─ 43 [ksoftirqd/7]
```

The `spawn_ksoftirqd` function starts this these threads. As we can see this function called as early initcall:

```
early_initcall(spawn_ksoftirqd);
```

Softirqs are determined statically at compile-time of the Linux kernel and the `open_softirq` function takes care of `softirq` initialization. The `open_softirq` function defined in the kernel/softirq.c:

```
void open_softirq(int nr, void (*action)(struct softirq_action *))
{
    softirq_vec[nr].action = action;
}
```

and as we can see this function uses two parameters:

- the index of the `softirq_vec` array;
- a pointer to the softirq function to be executed;

First of all let's look on the `softirq_vec` array:

```
static struct softirq_action softirq_vec[NR_SOFTIRQS] __cacheline_aligned_in_smp;
```

it defined in the same source code file. As we can see, the `softirq_vec` array may contain `NR_SOFTIRQS` or `10` types of `softirqs` that has type `softirq_action`. First of all about its elements. In the current version of the Linux kernel there are ten softirq vectors defined; two for tasklet processing, two for networking, two for the block layer, two for timers, and one each for the scheduler and read-copy-update processing. All of these kinds are represented by the following enum:

```
enum
{
        HI_SOFTIRQ=0,
        TIMER_SOFTIRQ,
        NET_TX_SOFTIRQ,
        NET_RX_SOFTIRQ,
        BLOCK_SOFTIRQ,
        BLOCK_IOPOLL_SOFTIRQ,
        TASKLET_SOFTIRQ,
        SCHED_SOFTIRQ,
        HRTIMER_SOFTIRQ,
        RCU_SOFTIRQ,
        NR_SOFTIRQS
};
```

All names of these kinds of softirqs are represented by the following array:

```
const char * const softirq_to_name[NR_SOFTIRQS] = {
        "HI", "TIMER", "NET_TX", "NET_RX", "BLOCK", "BLOCK_IOPOLL",
        "TASKLET", "SCHED", "HRTIMER", "RCU"
};
```

Or we can see it in the output of the `/proc/softirqs`:

```
~$ cat /proc/softirqs
                    CPU0       CPU1       CPU2       CPU3       CPU4       CPU5       CPU6       CPU7
          HI:          5          0          0          0          0          0          0          0
       TIMER:     332519     310498     289555     272913     282535     279467     282895     270979
      NET_TX:       2320          0          0          2          1          1          0          0
      NET_RX:     270221        225        338        281        311        262        430        265
```

```
        BLOCK:     134282         32         40         10         12          7          8          8
 BLOCK_IOPOLL:          0          0          0          0          0          0          0          0
      TASKLET:     196835          2          3          0          0          0          0          0
        SCHED:     161852     146745     129539     126064     127998     128014     120243     117391
      HRTIMER:          0          0          0          0          0          0          0          0
          RCU:     337707     289397     251874     239796     254377     254898     267497     256624
```

As we can see the `softirq_vec` array has `softirq_action` types. This is the main data structure related to the `softirq` mechanism, so all `softirqs` represented by the `softirq_action` structure. The `softirq_action` structure consists a single field only: an action pointer to the softirq function:

```
struct softirq_action
{
         void    (*action)(struct softirq_action *);
};
```

So, after this we can understand that the `open_softirq` function fills the `softirq_vec` array with the given `softirq_action` . The registered deferred interrupt (with the call of the `open_softirq` function) for it to be queued for execution, it should be activated by the call of the `raise_softirq` function. This function takes only one parameter -- a softirq index `nr` . Let's look on its implementation:

```
void raise_softirq(unsigned int nr)
{
        unsigned long flags;

        local_irq_save(flags);
        raise_softirq_irqoff(nr);
        local_irq_restore(flags);
}
```

Here we can see the call of the `raise_softirq_irqoff` function between the `local_irq_save` and the `local_irq_restore` macros. The `local_irq_save` defined in the include/linux/irqflags.h header file and saves the state of the IF flag of the eflags register and disables interrupts on the local processor. The `local_irq_restore` macro defined in the same header file and does the opposite thing: restores the `interrupt flag` and enables interrupts. We disable interrupts here because a `softirq` interrupt runs in the interrupt context and that one softirq (and no others) will be run.

The `raise_softirq_irqoff` function marks the softirq as deffered by setting the bit corresponding to the given index `nr` in the `softirq` bit mask (`__softirq_pending`) of the local processor. It does it with the help of the:

```
__raise_softirq_irqoff(nr);
```

macro. After this, it checks the result of the `in_interrupt` that returns `irq_count` value. We already saw the `irq_count` in the first part of this chapter and it is used to check if a CPU is already on an interrupt stack or not. We just exit from the `raise_softirq_irqoff` , restore `IF` flag and enable interrupts on the local processor, if we are in the interrupt context, otherwise we call the `wakeup_softirqd` :

```
if (!in_interrupt())
    wakeup_softirqd();
```

Where the `wakeup_softirqd` function activates the `ksoftirqd` kernel thread of the local processor:

```
static void wakeup_softirqd(void)
{
    struct task_struct *tsk = __this_cpu_read(ksoftirqd);

    if (tsk && tsk->state != TASK_RUNNING)
        wake_up_process(tsk);
```

```
}
```

Each `ksoftirqd` kernel thread runs the `run_ksoftirqd` function that checks existence of deferred interrupts and calls the `__do_softirq` function depending on the result of the check. This function reads the `__softirq_pending` softirq bit mask of the local processor and executes the deferrable functions corresponding to every bit set. During execution of a deferred function, new pending `softirqs` might occur. The main problem here that execution of the userspace code can be delayed for a long time while the `__do_softirq` function will handle deferred interrupts. For this purpose, it has the limit of the time when it must be finished:

```
unsigned long end = jiffies + MAX_SOFTIRQ_TIME;
...
...
...
restart:
while ((softirq_bit = ffs(pending))) {
    ...
    h->action(h);
    ...
}
...
...
...
pending = local_softirq_pending();
if (pending) {
    if (time_before(jiffies, end) && !need_resched() &&
        --max_restart)
            goto restart;
}
...
```

Checks of the existence of the deferred interrupts are performed periodically. There are several points where these checks occur. The main point is the call of the `do_IRQ` function defined in arch/x86/kernel/irq.c, which provides the main means for actual interrupt processing in the Linux kernel. When `do_IRQ` finishes handling an interrupt, it calls the `exiting_irq` function from the arch/x86/include/asm/apic.h that expands to the call of the `irq_exit` function. `irq_exit` checks for deferred interrupts and the current context and calls the `invoke_softirq` function:

```
if (!in_interrupt() && local_softirq_pending())
    invoke_softirq();
```

that also executes `__do_softirq`. To summarize, each `softirq` goes through the following stages:

- Registration of a `softirq` with the `open_softirq` function.
- Activation of a `softirq` by marking it as deferred with the `raise_softirq` function.
- After this, all marked `softirqs` will be triggered in the next time the Linux kernel schedules a round of executions of deferrable functions.
- And execution of the deferred functions that have the same type.

As I already wrote, the `softirqs` are statically allocated and it is a problem for a kernel module that can be loaded. The second concept that built on top of `softirq` -- the `tasklets` solves this problem.

Tasklets

If you read the source code of the Linux kernel that is related to the `softirq`, you notice that it is used very rarely. The preferable way to implement deferrable functions are `tasklets`. As I already wrote above the `tasklets` are built on top of the `softirq` concept and generally on top of two `softirqs`:

- `TASKLET_SOFTIRQ` ;
- `HI_SOFTIRQ` .

In short words, `tasklets` are `softirqs` that can be allocated and initialized at runtime and unlike `softirqs` , tasklets that have the same type cannot be run on multiple processors at a time. Ok, now we know a little bit about the `softirqs` , of course previous text does not cover all aspects about this, but now we can directly look on the code and to know more about the `softirqs` step by step on practice and to know about `tasklets` . Let's return back to the implementation of the `softirq_init` function that we talked about in the beginning of this part. This function is defined in the kernel/softirq.c source code file, let's look on its implementation:

```
void __init softirq_init(void)
{
        int cpu;

        for_each_possible_cpu(cpu) {
                per_cpu(tasklet_vec, cpu).tail =
                        &per_cpu(tasklet_vec, cpu).head;
                per_cpu(tasklet_hi_vec, cpu).tail =
                        &per_cpu(tasklet_hi_vec, cpu).head;
        }

        open_softirq(TASKLET_SOFTIRQ, tasklet_action);
        open_softirq(HI_SOFTIRQ, tasklet_hi_action);
}
```

We can see definition of the integer `cpu` variable at the beginning of the `softirq_init` function. Next we will use it as parameter for the `for_each_possible_cpu` macro that goes through the all possible processors in the system. If the `possible processor` is the new terminology for you, you can read more about it the CPU masks chapter. In short words, `possible cpus` is the set of processors that can be plugged in anytime during the life of that system boot. All `possible processors` stored in the `cpu_possible_bits` bitmap, you can find its definition in the kernel/cpu.c:

```
static DECLARE_BITMAP(cpu_possible_bits, CONFIG_NR_CPUS) __read_mostly;
...
...
...
const struct cpumask *const cpu_possible_mask = to_cpumask(cpu_possible_bits);
```

Ok, we defined the integer `cpu` variable and go through the all possible processors with the `for_each_possible_cpu` macro and makes initialization of the two following per-cpu variables:

- `tasklet_vec` ;
- `tasklet_hi_vec` ;

These two `per-cpu` variables defined in the same source code file as the `softirq_init` function and represent two `tasklet_head` structures:

```
static DEFINE_PER_CPU(struct tasklet_head, tasklet_vec);
static DEFINE_PER_CPU(struct tasklet_head, tasklet_hi_vec);
```

Where `tasklet_head` structure represents a list of `Tasklets` and contains two fields, head and tail:

```
struct tasklet_head {
        struct tasklet_struct *head;
        struct tasklet_struct **tail;
};
```

The `tasklet_struct` structure is defined in the include/linux/interrupt.h and represents the `Tasklet`. Previously we did not see this word in this book. Let's try to understand what the `tasklet` is. Actually, the tasklet is one of mechanisms to handle deferred interrupt. Let's look on the implementation of the `tasklet_struct` structure:

```
struct tasklet_struct
{
        struct tasklet_struct *next;
        unsigned long state;
        atomic_t count;
        void (*func)(unsigned long);
        unsigned long data;
};
```

As we can see this structure contains five fields, they are:

- Next tasklet in the scheduling queue;
- State of the tasklet;
- Represent current state of the tasklet, active or not;
- Main callback of the tasklet;
- Parameter of the callback.

In our case, we set only for initialize only two arrays of tasklets in the `softirq_init` function: the `tasklet_vec` and the `tasklet_hi_vec`. Tasklets and high-priority tasklets are stored in the `tasklet_vec` and `tasklet_hi_vec` arrays, respectively. So, we have initialized these arrays and now we can see two calls of the `open_softirq` function that is defined in the kernel/softirq.c source code file:

```
open_softirq(TASKLET_SOFTIRQ, tasklet_action);
open_softirq(HI_SOFTIRQ, tasklet_hi_action);
```

at the end of the `softirq_init` function. The main purpose of the `open_softirq` function is the initialization of `softirq`. Let's look on the implementation of the `open_softirq` function.

, in our case they are: `tasklet_action` and the `tasklet_hi_action` or the `softirq` function associated with the `HI_SOFTIRQ` softirq is named `tasklet_hi_action` and `softirq` function associated with the `TASKLET_SOFTIRQ` is named `tasklet_action`. The Linux kernel provides API for the manipulating of `tasklets`. First of all it is the `tasklet_init` function that takes `tasklet_struct`, function and parameter for it and initializes the given `tasklet_struct` with the given data:

```
void tasklet_init(struct tasklet_struct *t,
                  void (*func)(unsigned long), unsigned long data)
{
    t->next = NULL;
    t->state = 0;
    atomic_set(&t->count, 0);
    t->func = func;
    t->data = data;
}
```

There are additional methods to initialize a tasklet statically with the two following macros:

```
DECLARE_TASKLET(name, func, data);
DECLARE_TASKLET_DISABLED(name, func, data);
```

The Linux kernel provides three following functions to mark a tasklet as ready to run:

```
void tasklet_schedule(struct tasklet_struct *t);
void tasklet_hi_schedule(struct tasklet_struct *t);
void tasklet_hi_schedule_first(struct tasklet_struct *t);
```

The first function schedules a tasklet with the normal priority, the second with the high priority and the third out of turn. Implementation of the all of these three functions is similar, so we will consider only the first -- `tasklet_schedule`. Let's look on its implementation:

```
static inline void tasklet_schedule(struct tasklet_struct *t)
{
    if (!test_and_set_bit(TASKLET_STATE_SCHED, &t->state))
        __tasklet_schedule(t);
}

void __tasklet_schedule(struct tasklet_struct *t)
{
        unsigned long flags;

        local_irq_save(flags);
        t->next = NULL;
        *__this_cpu_read(tasklet_vec.tail) = t;
        __this_cpu_write(tasklet_vec.tail, &(t->next));
        raise_softirq_irqoff(TASKLET_SOFTIRQ);
        local_irq_restore(flags);
}
```

As we can see it checks and sets the state of the given tasklet to the `TASKLET_STATE_SCHED` and executes the `__tasklet_schedule` with the given tasklet. The `__tasklet_schedule` looks very similar to the `raise_softirq` function that we saw above. It saves the `interrupt flag` and disables interrupts at the beginning. After this, it updates `tasklet_vec` with the new tasklet and calls the `raise_softirq_irqoff` function that we saw above. When the Linux kernel scheduler will decide to run deferred functions, the `tasklet_action` function will be called for deferred functions which are associated with the `TASKLET_SOFTIRQ` and `tasklet_hi_action` for deferred functions which are associated with the `HI_SOFTIRQ`. These functions are very similar and there is only one difference between them -- `tasklet_action` uses `tasklet_vec` and `tasklet_hi_action` uses `tasklet_hi_vec`.

Let's look on the implementation of the `tasklet_action` function:

```
static void tasklet_action(struct softirq_action *a)
{
    local_irq_disable();
    list = __this_cpu_read(tasklet_vec.head);
    __this_cpu_write(tasklet_vec.head, NULL);
    __this_cpu_write(tasklet_vec.tail, this_cpu_ptr(&tasklet_vec.head));
    local_irq_enable();

    while (list) {
        if (tasklet_trylock(t)) {
            t->func(t->data);
            tasklet_unlock(t);
        }
        ...
        ...
        ...
    }
}
```

In the beginning of the `tasklet_action` function, we disable interrupts for the local processor with the help of the `local_irq_disable` macro (you can read about this macro in the second part of this chapter). In the next step, we take a head of the list that contains tasklets with normal priority and set this per-cpu list to `NULL` because all tasklets must be executed in a generally way. After this we enable interrupts for the local processor and go through the list of tasklets in the loop. In every iteration of the loop we call the `tasklet_trylock` function for the given tasklet that updates state of the given tasklet on `TASKLET_STATE_RUN`:

```
static inline int tasklet_trylock(struct tasklet_struct *t)
{
    return !test_and_set_bit(TASKLET_STATE_RUN, &(t)->state);
}
```

If this operation was successful we execute tasklet's action (it was set in the `tasklet_init`) and call the `tasklet_unlock` function that clears tasklet's `TASKLET_STATE_RUN` state.

In general, that's all about `tasklets` concept. Of course this does not cover full `tasklets` , but I think that it is a good point from where you can continue to learn this concept.

The `tasklets` are widely used concept in the Linux kernel, but as I wrote in the beginning of this part there is third mechanism for deferred functions -- `workqueue` . In the next paragraph we will see what it is.

Workqueues

The `workqueue` is another concept for handling deferred functions. It is similar to `tasklets` with some differences. Workqueue functions run in the context of a kernel process, but `tasklet` functions run in the software interrupt context. This means that `workqueue` functions must not be atomic as `tasklet` functions. Tasklets always run on the processor from which they were originally submitted. Workqueues work in the same way, but only by default. The `workqueue` concept represented by the:

```
struct worker_pool {
    spinlock_t              lock;
    int                     cpu;
    int                     node;
    int                     id;
    unsigned int            flags;

    struct list_head        worklist;
    int                     nr_workers;
...
...
...
```

structure that is defined in the kernel/workqueue.c source code file in the Linux kernel. I will not write the source code of this structure here, because it has quite a lot of fields, but we will consider some of those fields.

In its most basic form, the work queue subsystem is an interface for creating kernel threads to handle work that is queued from elsewhere. All of these kernel threads are called -- `worker threads` . The work queue are maintained by the `work_struct` that defined in the include/linux/workqueue.h. Let's look on this structure:

```
struct work_struct {
    atomic_long_t data;
    struct list_head entry;
    work_func_t func;
#ifdef CONFIG_LOCKDEP
    struct lockdep_map lockdep_map;
#endif
};
```

Here are two things that we are interested: `func` -- the function that will be scheduled by the `workqueue` and the `data` - parameter of this function. The Linux kernel provides special per-cpu threads that are called `kworker` :

```
systemd-cgls -k | grep kworker
├─    5 [kworker/0:0H]
├─   15 [kworker/1:0H]
├─   20 [kworker/2:0H]
```

```
├─   25 [kworker/3:0H]
├─   30 [kworker/4:0H]
...
...
...
```

This process can be used to schedule the deferred functions of the workqueues (as `ksoftirqd` for `softirqs`). Besides this we can create new separate worker thread for a `workqueue`. The Linux kernel provides following macros for the creation of workqueue:

```
#define DECLARE_WORK(n, f) \
    struct work_struct n = __WORK_INITIALIZER(n, f)
```

for static creation. It takes two parameters: name of the workqueue and the workqueue function. For creation of workqueue in runtime, we can use the:

```
#define INIT_WORK(_work, _func)       \
    __INIT_WORK((_work), (_func), 0)

#define __INIT_WORK(_work, _func, _onstack)                     \
    do {                                                        \
            __init_work((_work), _onstack);                     \
            (_work)->data = (atomic_long_t) WORK_DATA_INIT();   \
            INIT_LIST_HEAD(&(_work)->entry);                    \
             (_work)->func = (_func);                           \
    } while (0)
```

macro that takes `work_struct` structure that has to be created and the function to be scheduled in this workqueue. After a `work` was created with the one of these macros, we need to put it to the `workqueue`. We can do it with the help of the `queue_work` or the `queue_delayed_work` functions:

```
static inline bool queue_work(struct workqueue_struct *wq,
                              struct work_struct *work)
{
    return queue_work_on(WORK_CPU_UNBOUND, wq, work);
}
```

The `queue_work` function just calls the `queue_work_on` function that queue work on specific processor. Note that in our case we pass the `WORK_CPU_UNBOUND` to the `queue_work_on` function. It is a part of the `enum` that is defined in the include/linux/workqueue.h and represents workqueue which are not bound to any specific processor. The `queue_work_on` function tests and set the `WORK_STRUCT_PENDING_BIT` bit of the given `work` and executes the `__queue_work` function with the `workqueue` for the given processor and given `work`:

```
bool queue_work_on(int cpu, struct workqueue_struct *wq,
           struct work_struct *work)
{
    bool ret = false;
    ...
    if (!test_and_set_bit(WORK_STRUCT_PENDING_BIT, work_data_bits(work))) {
        __queue_work(cpu, wq, work);
        ret = true;
    }
    ...
    return ret;
}
```

The `__queue_work` function gets the `work pool`. Yes, the `work pool` not `workqueue`. Actually, all `works` are not placed in the `workqueue`, but to the `work pool` that is represented by the `worker_pool` structure in the Linux kernel. As you can see above, the `workqueue_struct` structure has the `pwqs` field which is list of `worker_pools`. When we create a `workqueue`, it stands out for each processor the `pool_workqueue`. Each `pool_workqueue` associated with `worker_pool`, which is allocated on the same processor and corresponds to the type of priority queue. Through them `workqueue` interacts with `worker_pool`. So in the `__queue_work` function we set the cpu to the current processor with the `raw_smp_processor_id` (you can find information about this macro in the fourth part of the Linux kernel initialization process chapter), getting the `pool_workqueue` for the given `workqueue_struct` and insert the given `work` to the given `workqueue`:

```
static void __queue_work(int cpu, struct workqueue_struct *wq,
                         struct work_struct *work)
{
...
...
...
if (req_cpu == WORK_CPU_UNBOUND)
    cpu = raw_smp_processor_id();

if (!(wq->flags & WQ_UNBOUND))
    pwq = per_cpu_ptr(wq->cpu_pwqs, cpu);
else
    pwq = unbound_pwq_by_node(wq, cpu_to_node(cpu));
...
...
...
insert_work(pwq, work, worklist, work_flags);
```

As we can create `works` and `workqueue`, we need to know when they are executed. As I already wrote, all `works` are executed by the kernel thread. When this kernel thread is scheduled, it starts to execute `works` from the given `workqueue`. Each worker thread executes a loop inside the `worker_thread` function. This thread makes many different things and part of these things are similar to what we saw before in this part. As it starts executing, it removes all `work_struct` or `works` from its `workqueue`.

That's all.

Conclusion

It is the end of the ninth part of the Interrupts and Interrupt Handling chapter and we continued to dive into external hardware interrupts in this part. In the previous part we saw initialization of the `IRQs` and main `irq_desc` structure. In this part we saw three concepts: the `softirq`, `tasklet` and `workqueue` that are used for the deferred functions.

The next part will be last part of the `Interrupts and Interrupt Handling` chapter and we will look on the real hardware driver and will try to learn how it works with the interrupts subsystem.

If you have any questions or suggestions, write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- initcall
- IF
- eflags
- CPU masks
- per-cpu

- Workqueue
- Previous part

Interrupts and Interrupt Handling. Part 10.

Last part

This is the tenth part of the chapter about interrupts and interrupt handling in the Linux kernel and in the previous part we saw a little about deferred interrupts and related concepts like `softirq`, `tasklet` and `workqeue`. In this part we will continue to dive into this theme and now it's time to look at real hardware driver.

Let's consider serial driver of the StrongARM** SA-110/21285 Evaluation Board board for example and will look how this driver requests an IRQ line, what happens when an interrupt is triggered and etc. The source code of this driver is placed in the drivers/tty/serial/21285.c source code file. Ok, we have source code, let's start.

Initialization of a kernel module

We will start to consider this driver as we usually did it with all new concepts that we saw in this book. We will start to consider it from the intialization. As you already may know, the Linux kernel provides two macros for initialization and finalization of a driver or a kernel module:

- `module_init`;
- `module_exit`.

And we can find usage of these macros in our driver source code:

```
module_init(serial21285_init);
module_exit(serial21285_exit);
```

The most part of device drivers can be compiled as a loadable kernel module or in another way they can be statically linked into the Linux kernel. In the first case initialization of a device driver will be produced via the `module_init` and `module_exit` macros that are defined in the include/linux/init.h:

```
#define module_init(initfn)                                     \
        static inline initcall_t __inittest(void)               \
        { return initfn; }                                      \
        int init_module(void) __attribute__((alias(#initfn)));

#define module_exit(exitfn)                                     \
        static inline exitcall_t __exittest(void)               \
        { return exitfn; }                                      \
        void cleanup_module(void) __attribute__((alias(#exitfn)));
```

and will be called by the initcall functions:

- `early_initcall`
- `pure_initcall`
- `core_initcall`
- `postcore_initcall`
- `arch_initcall`
- `subsys_initcall`
- `fs_initcall`
- `rootfs_initcall`
- `device_initcall`

- `late_initcall`

that are called in the `do_initcalls` from the init/main.c. Otherwise, if a device driver is statically linked into the Linux kernel, implementation of these macros will be following:

```
#define module_init(x)  __initcall(x);
#define module_exit(x)  __exitcall(x);
```

In this way implementation of module loading placed in the kernel/module.c source code file and initialization occurs in the `do_init_module` function. We will not dive into details about loadable modules in this chapter, but will see it in the special chapter that will describe Linux kernel modules. Ok, the `module_init` macro takes one parameter - the `serial21285_init` in our case. As we can understand from function's name, this function does stuff related to the driver initialization. Let's look at it:

```
static int __init serial21285_init(void)
{
    int ret;

    printk(KERN_INFO "Serial: 21285 driver\n");

    serial21285_setup_ports();

    ret = uart_register_driver(&serial21285_reg);
    if (ret == 0)
        uart_add_one_port(&serial21285_reg, &serial21285_port);

    return ret;
}
```

As we can see, first of all it prints information about the driver to the kernel buffer and the call of the `serial21285_setup_ports` function. This function setups the base uart clock of the `serial21285_port` device:

```
unsigned int mem_fclk_21285 = 50000000;

static void serial21285_setup_ports(void)
{
    serial21285_port.uartclk = mem_fclk_21285 / 4;
}
```

Here the `serial21285` is the structure that describes `uart` driver:

```
static struct uart_driver serial21285_reg = {
    .owner            = THIS_MODULE,
    .driver_name    = "ttyFB",
    .dev_name        = "ttyFB",
    .major            = SERIAL_21285_MAJOR,
    .minor            = SERIAL_21285_MINOR,
    .nr                = 1,
    .cons            = SERIAL_21285_CONSOLE,
};
```

If the driver registered successfully we attach the driver-defined port `serial21285_port` structure with the `uart_add_one_port` function from the drivers/tty/serial/serial_core.c source code file and return from the `serial21285_init` function:

```
if (ret == 0)
    uart_add_one_port(&serial21285_reg, &serial21285_port);

return ret;
```

That's all. Our driver is initialized. When an `uart` port will be opened with the call of the `uart_open` function from the drivers/tty/serial/serial_core.c, it will call the `uart_startup` function to start up the serial port. This function will call the `startup` function that is part of the `uart_ops` structure. Each `uart` driver has the definition of this structure, in our case it is:

```
static struct uart_ops serial21285_ops = {
    ...
    .startup    = serial21285_startup,
    ...
}
```

`serial21285` structure. As we can see the `.strartup` field references on the `serial21285_startup` function. Implementation of this function is very interesting for us, because it is related to the interrupts and interrupt handling.

Requesting irq line

Let's look at the implementation of the `serial21285` function:

```
static int serial21285_startup(struct uart_port *port)
{
    int ret;

    tx_enabled(port) = 1;
    rx_enabled(port) = 1;

    ret = request_irq(IRQ_CONRX, serial21285_rx_chars, 0,
              serial21285_name, port);
    if (ret == 0) {
        ret = request_irq(IRQ_CONTX, serial21285_tx_chars, 0,
                  serial21285_name, port);
        if (ret)
            free_irq(IRQ_CONRX, port);
    }

    return ret;
}
```

First of all about `TX` and `RX`. A serial bus of a device consists of just two wires: one for sending data and another for receiving. As such, serial devices should have two serial pins: the receiver - `RX`, and the transmitter - `TX`. With the call of first two macros: `tx_enabled` and `rx_enabled`, we enable these wires. The following part of these function is the greatest interest for us. Note on `request_irq` functions. This function registers an interrupt handler and enables a given interrupt line. Let's look at the implementation of this function and get into the details. This function defined in the include/linux/interrupt.h header file and looks as:

```
static inline int __must_check
request_irq(unsigned int irq, irq_handler_t handler, unsigned long flags,
            const char *name, void *dev)
{
        return request_threaded_irq(irq, handler, NULL, flags, name, dev);
}
```

As we can see, the `request_irq` function takes five parameters:

- `irq` - the interrupt number that being requested;
- `handler` - the pointer to the interrupt handler;
- `flags` - the bitmask options;
- `name` - the name of the owner of an interrupt;
- `dev` - the pointer used for shared interrupt lines;

Now let's look at the calls of the `request_irq` functions in our example. As we can see the first parameter is `IRQ_CONRX`. We know that it is number of the interrupt, but what is it `CONRX`? This macro defined in the arch/arm/mach-footbridge/include/mach/irqs.h header file. We can find the full list of interrupts that the `21285` board can generate. Note that in the second call of the `request_irq` function we pass the `IRQ_CONTX` interrupt number. Both these interrupts will handle `RX` and `TX` event in our driver. Implementation of these macros is easy:

```
#define IRQ_CONRX               _DC21285_IRQ(0)
#define IRQ_CONTX               _DC21285_IRQ(1)
...
...
...
#define _DC21285_IRQ(x)         (16 + (x))
```

The ISA IRQs on this board are from `0` to `15`, so, our interrupts will have first two numbers: `16` and `17`. Second parameters for two calls of the `request_irq` functions are `serial21285_rx_chars` and `serial21285_tx_chars`. These functions will be called when an `RX` or `TX` interrupt occurred. We will not dive in this part into details of these functions, because this chapter covers the interrupts and interrupts handling but not device and drivers. The next parameter - `flags` and as we can see, it is zero in both calls of the `request_irq` function. All acceptable flags are defined as `IRQF_*` macros in the include/linux/interrupt.h. Some of it:

- `IRQF_SHARED` - allows sharing the irq among several devices;
- `IRQF_PERCPU` - an interrupt is per cpu;
- `IRQF_NO_THREAD` - an interrupt cannot be threaded;
- `IRQF_NOBALANCING` - excludes this interrupt from irq balancing;
- `IRQF_IRQPOLL` - an interrupt is used for polling;
- and etc.

In our case we pass `0`, so it will be `IRQF_TRIGGER_NONE`. This flag means that it does not imply any kind of edge or level triggered interrupt behaviour. To the fourth parameter (`name`), we pass the `serial21285_name` that defined as:

```
static const char serial21285_name[] = "Footbridge UART";
```

and will be displayed in the output of the `/proc/interrupts`. And in the last parameter we pass the pointer to the our main `uart_port` structure. Now we know a little about `request_irq` function and its parameters, let's look at its implemenetation. As we can see above, the `request_irq` function just makes a call of the `request_threaded_irq` function inside. The `request_threaded_irq` function defined in the kernel/irq/manage.c source code file and allocates a given interrupt line. If we will look at this function, it starts from the definition of the `irqaction` and the `irq_desc`:

```
int request_threaded_irq(unsigned int irq, irq_handler_t handler,
                         irq_handler_t thread_fn, unsigned long irqflags,
                         const char *devname, void *dev_id)
{
        struct irqaction *action;
        struct irq_desc *desc;
        int retval;
        ...
        ...
        ...
}
```

We arelady saw the `irqaction` and the `irq_desc` structures in this chapter. The first structure represents per interrupt action descriptor and contains pointers to the interrupt handler, name of the device, interrupt number, etc. The second structure represents a descriptor of an interrupt and contains pointer to the `irqaction`, interrupt flags, etc. Note that the `request_threaded_irq` function called by the `request_irq` with the additioanal parameter: `irq_handler_t thread_fn` . If this parameter is not `NULL` , the `irq` thread will be created and the given `irq` handler will be executed in this thread. In the next step we need to make following checks:

```
if (((irqflags & IRQF_SHARED) && !dev_id) ||
            (!(irqflags & IRQF_SHARED) && (irqflags & IRQF_COND_SUSPEND)) ||
            ((irqflags & IRQF_NO_SUSPEND) && (irqflags & IRQF_COND_SUSPEND)))
               return -EINVAL;
```

First of all we check that real `dev_id` is passed for the shared interrupt and the `IRQF_COND_SUSPEND` only makes sense for shared interrupts. Otherwise we exit from this function with the `-EINVAL` error. After this we convert the given `irq` number to the `irq` descriptor wit the help of the `irq_to_desc` function that defined in the kernel/irq/irqdesc.c source code file and exit from this function with the `-EINVAL` error if it was not successful:

```
desc = irq_to_desc(irq);
if (!desc)
    return -EINVAL;
```

The `irq_to_desc` function checks that given `irq` number is less than maximum number of IRQs and returns the irq descriptor where the `irq` number is offset from the `irq_desc` array:

```
struct irq_desc *irq_to_desc(unsigned int irq)
{
        return (irq < NR_IRQS) ? irq_desc + irq : NULL;
}
```

As we have converted `irq` number to the `irq` descriptor we make the check the status of the descriptor that an interrupt can be requested:

```
if (!irq_settings_can_request(desc) || WARN_ON(irq_settings_is_per_cpu_devid(desc)))
    return -EINVAL;
```

and exit with the `-EINVAL` in othre way. After this we check the given interrupt handler. If it was not passed to the `request_irq` function, we check the `thread_fn` . If both handlers are `NULL` , we return with the `-EINVAL` . If an interrupt handler was not passed to the `request_irq` function, but the `thread_fn` is not null, we set handler to the `irq_default_primary_handler` :

```
if (!handler) {
    if (!thread_fn)
        return -EINVAL;
    handler = irq_default_primary_handler;
}
```

In the next step we allocate memory for our `irqaction` with the `kzalloc` function and return from the function if this operation was not successful:

```
action = kzalloc(sizeof(struct irqaction), GFP_KERNEL);
if (!action)
    return -ENOMEM;
```

More about `kzalloc` will be in the separate chapter about memory management in the Linux kernel. As we allocated space for the `irqaction` , we start to initialize this structure with the values of interrupt handler, interrupt flags, device name, etc:

```
action->handler = handler;
action->thread_fn = thread_fn;
action->flags = irqflags;
```

```
action->name = devname;
action->dev_id = dev_id;
```

In the end of the `request_threaded_irq` function we call the `__setup_irq` function from the kernel/irq/manage.c and registers a given `irqaction`. Release memory for the `irqaction` and return:

```
chip_bus_lock(desc);
retval = __setup_irq(irq, desc, action);
chip_bus_sync_unlock(desc);

if (retval)
    kfree(action);

return retval;
```

Note that the call of the `__setup_irq` function is placed between the `chip_bus_lock` and the `chip_bus_sync_unlock` functions. These functions locl/unlock access to slow bus (like i2c) chips. Now let's look at the implementation of the `__setup_irq` function. In the beginning of the `__setup_irq` function we can see a couple of different checks. First of all we check that the given interrupt descriptor is not `NULL`, `irqchip` is not `NULL` and that given interrupt descriptor module owner is not `NULL`. After this we check is interrupt nest into another interrupt thread or not, and if it is nested we replace the `irq_default_primary_handler` with the `irq_nested_primary_handler`.

In the next step we create an irq handler thread with the `kthread_create` function, if the given interrupt is not nested and the `thread_fn` is not `NULL`:

```
if (new->thread_fn && !nested) {
    struct task_struct *t;
    t = kthread_create(irq_thread, new, "irq/%d-%s", irq, new->name);
    ...
}
```

And fill the rest of the given interrupt descriptor fields in the end. So, our `16` and `17` interrupt request lines are registered and the `serial21285_rx_chars` and `serial21285_tx_chars` functions will be invoked when an interrupt controller will get event releated to these interrupts. Now let's look at what happens when an interrupt occurs.

Prepare to handle an interrupt

In the previous paragraph we saw the requesting of the irq line for the given interrupt descriptor and registration of the `irqaction` structure for the given interrupt. We already know that when an interrupt event occurs, an interrupt controller notifies the processor about this event and processor tries to find appropriate interrupt gate for this interrupt. If you have read the eighth part of this chapter, you may remember the `native_init_IRQ` function. This function makes initialization of the local APIC. The following part of this function is the most interesting part for us right now:

```
for_each_clear_bit_from(i, used_vectors, first_system_vector) {
    set_intr_gate(i, irq_entries_start +
        8 * (i - FIRST_EXTERNAL_VECTOR));
}
```

Here we iterate over all the cleared bit of the `used_vectors` bitmap starting at `first_system_vector` that is:

```
int first_system_vector = FIRST_SYSTEM_VECTOR; // 0xef
```

and set interrupt gates with the `i` vector number and the `irq_entries_start + 8 * (i - FIRST_EXTERNAL_VECTOR)` start address. Only one things is unclear here - the `irq_entries_start`. This symbol defined in the arch/x86/entry/entry_64.S assembly file and provides `irq` entries. Let's look at it:

```
    .align 8
ENTRY(irq_entries_start)
    vector=FIRST_EXTERNAL_VECTOR
    .rept (FIRST_SYSTEM_VECTOR - FIRST_EXTERNAL_VECTOR)
    pushq    $(~vector+0x80)
    vector=vector+1
    jmp    common_interrupt
    .align    8
    .endr
END(irq_entries_start)
```

Here we can see the GNU assembler `.rept` instruction which repeats the sequence of lines that are before `.endr` - `FIRST_SYSTEM_VECTOR - FIRST_EXTERNAL_VECTOR` times. As we already know, the `FIRST_SYSTEM_VECTOR` is `0xef`, and the `FIRST_EXTERNAL_VECTOR` is equal to `0x20`. So, it will work:

```
>>> 0xef - 0x20
207
```

times. In the body of the `.rept` instruction we push entry stubs on the stack (note that we use negative numbers for the interrupt vector numbers, because positive numbers already reserved to identify system calls), increase the `vector` variable and jump on the `common_interrupt` label. In the `common_interrupt` we adjust vector number on the stack and execute `interrupt` number with the `do_IRQ` parameter:

```
common_interrupt:
    addq    $-0x80, (%rsp)
    interrupt do_IRQ
```

The macro `interrupt` defined in the same source code file and saves general purpose registers on the stack, change the userspace `gs` on the kernel with the `SWAPGS` assembler instruction if need, increase per-cpu - `irq_count` variable that shows that we are in interrupt and call the `do_IRQ` function. This function defined in the arch/x86/kernel/irq.c source code file and handles our device interrupt. Let's look at this function. The `do_IRQ` function takes one parameter - `pt_regs` structure that stores values of the userspace registers:n

```
__visible unsigned int __irq_entry do_IRQ(struct pt_regs *regs)
{
    struct pt_regs *old_regs = set_irq_regs(regs);
    unsigned vector = ~regs->orig_ax;
    unsigned irq;

    irq_enter();
    exit_idle();
    ...
    ...
    ...
}
```

At the beginning of this function we can see call of the `set_irq_regs` function that returns saved `per-cpu` irq register pointer and the calls of the `irq_enter` and `exit_idle` functions. The first function `irq_enter` enters to an interrupt context with the updating `__preempt_count` variable and the second function - `exit_idle` checks that current process is `idle` with pid - `0` and notify the `idle_notifier` with the `IDLE_END`.

In the next step we read the `irq` for the current cpu and call the `handle_irq` function:

```
irq = __this_cpu_read(vector_irq[vector]);

if (!handle_irq(irq, regs)) {
    ...
    ...
    ...
}
...
...
...
```

The `handle_irq` function defined in the arch/x86/kernel/irq_64.c source code file, checks the given interrupt descriptor and call the `generic_handle_irq_desc` :

```
desc = irq_to_desc(irq);
    if (unlikely(!desc))
        return false;
generic_handle_irq_desc(irq, desc);
```

Where the `generic_handle_irq_desc` calls the interrupt handler:

```
static inline void generic_handle_irq_desc(unsigned int irq, struct irq_desc *desc)
{
       desc->handle_irq(irq, desc);
}
```

But stop... What is it `handle_irq` and why do we call our interrupt handler from the interrupt descriptor when we know that `irqaction` points to the actual interrupt handler? Actually the `irq_desc->handle_irq` is a high-level API for the calling interrupt handler routine. It setups during initialization of the device tree and APIC initialization. The kernel selects correct function and call chain of the `irq->action(s)` there. In this way, the `serial21285_tx_chars` or the `serial21285_rx_chars` function will be executed after an interrupt will occur.

In the end of the `do_IRQ` function we call the `irq_exit` function that will exit from the interrupt context, the `set_irq_regs` with the old userspace registers and return:

```
irq_exit();
set_irq_regs(old_regs);
return 1;
```

We already know that when an `IRQ` finishes its work, deferred interrupts will be executed if they exist.

Exit from interrupt

Ok, the interrupt handler finished its execution and now we must return from the interrupt. When the work of the `do_IRQ` function will be finsihed, we will return back to the assembler code in the arch/x86/entry/entry_64.S to the `ret_from_intr` label. First of all we disable interrupts with the `DISABLE_INTERRUPTS` macro that expands to the `cli` instruction and decreases value of the `irq_count` per-cpu variable. Remember, this variable had value - `1` , when we were in interrupt context:

```
DISABLE_INTERRUPTS(CLBR_NONE)
TRACE_IRQS_OFF
decl    PER_CPU_VAR(irq_count)
```

In the last step we check the previous context (user or kernel), restore it in a correct way and exit from an interrupt with the:

```
INTERRUPT_RETURN
```

where the `INTERRUPT_RETURN` macro is:

```
#define INTERRUPT_RETURN    jmp native_iret
```

and

```
ENTRY(native_iret)

.global native_irq_return_iret
native_irq_return_iret:
    iretq
```

That's all.

Conclusion

It is the end of the tenth part of the Interrupts and Interrupt Handling chapter and as you have read in the beginning of this part - it is the last part of this chapter. This chapter started from the explanation of the theory of interrupts and we have learned what is it interrupt and kinds of interrupts, then we saw exceptions and handling of this kind of interrupts, deferred interrupts and finally we looked on the hardware interrupts and the handling of theirs in this part. Of course, this part and even this chapter does not cover full aspects of interrupts and interrupt handling in the Linux kernel. It is not realistic to do this. At least for me. It was the big part, I don't know how about you, but it was really big for me. This theme is much bigger than this chapter and I am not sure that somewhere there is a book that covers it. We have missed many part and aspects of interrupts and interrupt handling, but I think it will be good point to dive in the kernel code related to the interrupts and interrupts handling.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- Serial driver documentation
- StrongARM** SA-110/21285 Evaluation Board
- IRQ
- module
- initcall
- uart
- ISA
- memory managementd
- i2c
- APIC
- GNU assembler
- Processor register
- per-cpu
- pid
- device tree
- system calls
- Previous part

System calls

This chapter describes the `system call` concept in the linux kernel.

* Introduction to system call concept - this part is introduction to the `system call` concept in the Linux kernel.
* How the Linux kernel handles a system call - this part describes how the Linux kernel handles a system call from a userspace application.
* vsyscall and vDSO - third part describes `vsyscall` and `vDSO` concepts.
* How the Linux kernel runs a program - this part describes startup process of a program.
* Implementation of the open system call - this part describes implementation of the open system call.
* Limits on resources in Linux - this part describes implementation of the getrlimit/setrlimit system calls.

System calls in the Linux kernel. Part 1.

Introduction

This post opens up a new chapter in linux-insides book, and as you may understand from the title, this chapter will be devoted to the System call concept in the Linux kernel. The choice of topic for this chapter is not accidental. In the previous chapter we saw interrupts and interrupt handling. The concept of system calls is very similar to that of interrupts. This is because the most common way to implement system calls is as software interrupts. We will see many different aspects that are related to the system call concept. For example, we will learn what's happening when a system call occurs from userspace. We will see an implementation of a couple system call handlers in the Linux kernel, VDSO and vsyscall concepts and many many more.

Before we dive into Linux system call implementation, it is good to know some theory about system calls. Let's do it in the following paragraph.

System call. What is it?

A system call is just a userspace request of a kernel service. Yes, the operating system kernel provides many services. When your program wants to write to or read from a file, start to listen for connections on a socket, delete or create directory, or even to finish its work, a program uses a system call. In other words, a system call is just a C kernel space function that user space programs call to handle some request.

The Linux kernel provides a set of these functions and each architecture provides its own set. For example: the x86_64 provides 322 system calls and the x86 provides 358 different system calls. Ok, a system call is just a function. Let's look on a simple `Hello world` example that's written in the assembly programming language:

```
.data

msg:
    .ascii "Hello, world!\n"
    len = . - msg

.text
    .global _start

_start:
    movq  $1, %rax
    movq  $1, %rdi
    movq  $msg, %rsi
    movq  $len, %rdx
    syscall

    movq  $60, %rax
    xorq  %rdi, %rdi
    syscall
```

We can compile the above with the following commands:

```
$ gcc -c test.S
$ ld -o test test.o
```

and run it as follows:

```
./test
Hello, world!
```

Ok, what do we see here? This simple code represents `Hello world` assembly program for the Linux `x86_64` architecture. We can see two sections here:

* `.data`
* `.text`

The first section - `.data` stores initialized data of our program (`Hello world` string and its length in our case). The second section - `.text` contains the code of our program. We can split the code of our program into two parts: first part will be before the first `syscall` instruction and the second part will be between first and second `syscall` instructions. First of all what does the `syscall` instruction do in our code and generally? As we can read in the 64-ia-32-architectures-software-developer-vol-2b-manual:

```
SYSCALL invokes an OS system-call handler at privilege level 0. It does so by
loading RIP from the IA32_LSTAR MSR (after saving the address of the instruction
following SYSCALL into RCX). (The WRMSR instruction ensures that the
IA32_LSTAR MSR always contain a canonical address.)
...
...
...
SYSCALL loads the CS and SS selectors with values derived from bits 47:32 of the
IA32_STAR MSR. However, the CS and SS descriptor caches are not loaded from the
descriptors (in GDT or LDT) referenced by those selectors.

Instead, the descriptor caches are loaded with fixed values. It is the respon-
sibility of OS software to ensure that the descriptors (in GDT or LDT) referenced
by those selector values correspond to the fixed values loaded into the descriptor
caches; the SYSCALL instruction does not ensure this correspondence.
```

To summarize, the `syscall` instruction jumps to the address stored in the `MSR_LSTAR` Model specific register (Long system target address register). The kernel is responsible for providing its own custom function for handling syscalls as well as writing the address of this handler function to the `MSR_LSTAR` register upon system startup. The custom function is `entry_SYSCALL_64` , which is defined in arch/x86/entry/entry_64.S. The address of this syscall handling function is written to the `MSR_LSTAR` register during startup in arch/x86/kernel/cpu/common.c.

```
wrmsrl(MSR_LSTAR, entry_SYSCALL_64);
```

So, the `syscall` instruction invokes a handler of a given system call. But how does it know which handler to call? Actually it gets this information from the general purpose registers. As you can see in the system call table, each system call has a unique number. In our example the first system call is `write` , which writes data to the given file. Let's look in the system call table and try to find the `write` system call. As we can see, the write system call has number `1` . We pass the number of this system call through the `rax` register in our example. The next general purpose registers: `%rdi` , `%rsi` , and `%rdx` take the three parameters of the `write` syscall. In our case, they are:

* File descriptor (`1` is stdout in our case)
* Pointer to our string
* Size of data

Yes, you heard right. Parameters for a system call. As I already wrote above, a system call is a just `C` function in the kernel space. In our case first system call is write. This system call defined in the fs/read_write.c source code file and looks like:

```
SYSCALL_DEFINE3(write, unsigned int, fd, const char __user *, buf,
        size_t, count)
{
    ...
    ...
    ...
}
```

Or in other words:

```
ssize_t write(int fd, const void *buf, size_t nbytes);
```

Don't worry about the `SYSCALL_DEFINE3` macro for now, we'll come back to it.

The second part of our example is the same, but we call another system call. In this case we call the exit system call. This system call gets only one parameter:

- Return value

and handles the way our program exits. We can pass the program name of our program to the strace util and we will see our system calls:

```
$ strace test
execve("./test", ["./test"], [/* 62 vars */]) = 0
write(1, "Hello, world!\n", 14Hello, world!
)          = 14
_exit(0)                                = ?

+++ exited with 0 +++
```

In the first line of the `strace` output, we can see the execve system call that executes our program, and the second and third are system calls that we have used in our program: `write` and `exit`. Note that we pass the parameter through the general purpose registers in our example. The order of the registers is not accidental. The order of the registers is defined by the following agreement - x86-64 calling conventions. This, and the other agreement for the `x86_64` architecture are explained in the special document - System V Application Binary Interface. PDF. In a general way, argument(s) of a function are placed either in registers or pushed on the stack. The right order is:

- `rdi`
- `rsi`
- `rdx`
- `rcx`
- `r8`
- `r9`

for the first six parameters of a function. If a function has more than six arguments, the remaining parameters will be placed on the stack.

We do not use system calls in our code directly, but our program uses them when we want to print something, check access to a file or just write or read something to it.

For example:

```
#include <stdio.h>

int main(int argc, char **argv)
{
   FILE *fp;
   char buff[255];

   fp = fopen("test.txt", "r");
   fgets(buff, 255, fp);
   printf("%s\n", buff);
   fclose(fp);

   return 0;
}
```

There are no `fopen` , `fgets` , `printf` , and `fclose` system calls in the Linux kernel, but `open` , `read` , `write` , and `close` instead. I think you know that `fopen` , `fgets` , `printf` , and `fclose` are defined in the `C` standard library. Actually, these functions are just wrappers for the system calls. We do not call system calls directly in our code, but instead use these wrapper functions from the standard library. The main reason of this is simple: a system call must be performed quickly, very quickly. As a system call must be quick, it must be small. The standard library takes responsibility to perform system calls with the correct parameters and makes different checks before it will call the given system call. Let's compile our program with the following command:

```
$ gcc test.c -o test
```

and examine it with the ltrace util:

```
$ ltrace ./test
__libc_start_main([ "./test" ] <unfinished ...>
fopen("test.txt", "r")                                              = 0x602010
fgets("Hello World!\n", 255, 0x602010)                              = 0x7ffd2745e700
puts("Hello World!\n"Hello World!

)                                                                   = 14
fclose(0x602010)                                                    = 0
+++ exited (status 0) +++
```

The `ltrace` util displays a set of userspace calls of a program. The `fopen` function opens the given text file, the `fgets` function reads file content to the `buf` buffer, the `puts` function prints the buffer to `stdout` , and the `fclose` function closes the file given by the file descriptor. And as I already wrote, all of these functions call an appropriate system call. For example, `puts` calls the `write` system call inside, we can see it if we will add `-S` option to the `ltrace` program:

```
write@SYS(1, "Hello World!\n\n", 14) = 14
```

Yes, system calls are ubiquitous. Each program needs to open/write/read files and network connections, allocate memory, and many other things that can be provided only by the kernel. The proc file system contains special files in a format: `/proc/pid/systemcall` that exposes the system call number and argument registers for the system call currently being executed by the process. For example, pid 1 is systemd for me:

```
$ sudo cat /proc/1/comm
systemd

$ sudo cat /proc/1/syscall
232 0x4 0x7ffdf82e11b0 0x1f 0xffffffff 0x100 0x7ffdf82e11bf 0x7ffdf82e11a0 0x7f9114681193
```

the system call with number - `232` which is epoll_wait system call that waits for an I/O event on an epoll file descriptor. Or for example `emacs` editor where I'm writing this part:

```
$ ps ax | grep emacs
2093 ?        Sl     2:40 emacs

$ sudo cat /proc/2093/comm
emacs

$ sudo cat /proc/2093/syscall
270 0xf 0x7fff068a5a90 0x7fff068a5b10 0x0 0x7fff068a59c0 0x7fff068a59d0 0x7fff068a59b0 0x7f777dd8813c
```

the system call with the number `270` which is sys_pselect6 system call that allows `emacs` to monitor multiple file descriptors.

Now we know a little about system call, what is it and why we need in it. So let's look at the `write` system call that our program used.

Implementation of write system call

Let's look at the implementation of this system call directly in the source code of the Linux kernel. As we already know, the `write` system call is defined in the fs/read_write.c source code file and looks like this:

```
SYSCALL_DEFINE3(write, unsigned int, fd, const char __user *, buf,
        size_t, count)
{
    struct fd f = fdget_pos(fd);
    ssize_t ret = -EBADF;

    if (f.file) {
        loff_t pos = file_pos_read(f.file);
        ret = vfs_write(f.file, buf, count, &pos);
        if (ret >= 0)
            file_pos_write(f.file, pos);
        fdput_pos(f);
    }

    return ret;
}
```

First of all, the `SYSCALL_DEFINE3` macro is defined in the include/linux/syscalls.h header file and expands to the definition of the `sys_name(...)` function. Let's look at this macro:

```
#define SYSCALL_DEFINE3(name, ...) SYSCALL_DEFINEx(3, _##name, __VA_ARGS__)

#define SYSCALL_DEFINEx(x, sname, ...)                \
        SYSCALL_METADATA(sname, x, __VA_ARGS__)       \
        __SYSCALL_DEFINEx(x, sname, __VA_ARGS__)
```

As we can see the `SYSCALL_DEFINE3` macro takes `name` parameter which will represent name of a system call and variadic number of parameters. This macro just expands to the `SYSCALL_DEFINEx` macro that takes the number of the parameters the given system call, the `_##name` stub for the future name of the system call (more about tokens concatenation with the `##` you can read in the documentation of gcc). Next we can see the `SYSCALL_DEFINEx` macro. This macro expands to the two following macros:

- `SYSCALL_METADATA` ;
- `__SYSCALL_DEFINEx` .

Implementation of the first macro `SYSCALL_METADATA` depends on the `CONFIG_FTRACE_SYSCALLS` kernel configuration option. As we can understand from the name of this option, it allows to enable tracer to catch the syscall entry and exit events. If this kernel configuration option is enabled, the `SYSCALL_METADATA` macro executes initialization of the `syscall_metadata` structure that defined in the include/trace/syscall.h header file and contains different useful fields as name of a system call, number of a system call in the system call table, number of parameters of a system call, list of parameter types and etc:

```
#define SYSCALL_METADATA(sname, nb, ...)                             \
    ...                                                              \
    ...                                                              \
    ...                                                              \
    struct syscall_metadata __used                                   \
              __syscall_meta_##sname = {                             \
                    .name           = "sys"#sname,                   \
                    .syscall_nr     = -1,                            \
                    .nb_args        = nb,                            \
                    .types          = nb ? types_##sname : NULL,     \
                    .args           = nb ? args_##sname : NULL,      \
                    .enter_event    = &event_enter_##sname,          \
                    .exit_event     = &event_exit_##sname,           \
                    .enter_fields   = LIST_HEAD_INIT(__syscall_meta_##sname.enter_fields), \
             };                                                                            \
```

```
    static struct syscall_metadata __used                          \
              __attribute__((section("__syscalls_metadata")))      \
             *__p_syscall_meta_##sname = &__syscall_meta_##sname;
```

If the `CONFIG_FTRACE_SYSCALLS` kernel option is not enabled during kernel configuration, the `SYSCALL_METADATA` macro expands to an empty string:

```
#define SYSCALL_METADATA(sname, nb, ...)
```

The second macro `__SYSCALL_DEFINEx` expands to the definition of the five following functions:

```
#define __SYSCALL_DEFINEx(x, name, ...)                                 \
        asmlinkage long sys##name(__MAP(x,__SC_DECL,__VA_ARGS__))       \
                __attribute__((alias(__stringify(SyS##name))));         \
                                                                        \
        static inline long SYSC##name(__MAP(x,__SC_DECL,__VA_ARGS__));  \
                                                                        \
        asmlinkage long SyS##name(__MAP(x,__SC_LONG,__VA_ARGS__));      \
                                                                        \
        asmlinkage long SyS##name(__MAP(x,__SC_LONG,__VA_ARGS__))       \
        {                                                               \
                long ret = SYSC##name(__MAP(x,__SC_CAST,__VA_ARGS__));  \
                __MAP(x,__SC_TEST,__VA_ARGS__);                         \
                __PROTECT(x, ret,__MAP(x,__SC_ARGS,__VA_ARGS__));       \
                return ret;                                             \
        }                                                               \
                                                                        \
        static inline long SYSC##name(__MAP(x,__SC_DECL,__VA_ARGS__))
```

The first `sys##name` is definition of the syscall handler function with the given name - `sys_system_call_name`. The `__SC_DECL` macro takes the `__VA_ARGS__` and combines call input parameter system type and the parameter name, because the macro definition is unable to determine the parameter types. And the `__MAP` macro applies `__SC_DECL` macro to the `__VA_ARGS__` arguments. The other functions that are generated by the `__SYSCALL_DEFINEx` macro are need to protect from the CVE-2009-0029 and we will not dive into details about this here. Ok, as result of the `SYSCALL_DEFINE3` macro, we will have:

```
asmlinkage long sys_write(unsigned int fd, const char __user * buf, size_t count);
```

Now we know a little about the system call's definition and we can go back to the implementation of the `write` system call. Let's look on the implementation of this system call again:

```
SYSCALL_DEFINE3(write, unsigned int, fd, const char __user *, buf,
        size_t, count)
{
    struct fd f = fdget_pos(fd);
    ssize_t ret = -EBADF;

    if (f.file) {
        loff_t pos = file_pos_read(f.file);
        ret = vfs_write(f.file, buf, count, &pos);
        if (ret >= 0)
            file_pos_write(f.file, pos);
        fdput_pos(f);
    }

    return ret;
}
```

As we already know and can see from the code, it takes three arguments:

* `fd` - file descriptor;
* `buf` - buffer to write;
* `count` - length of buffer to write.

and writes data from a buffer declared by the user to a given device or a file. Note that the second parameter `buf` , defined with the `__user` attribute. The main purpose of this attribute is for checking the Linux kernel code with the sparse util. It is defined in the include/linux/compiler.h header file and depends on the `__CHECKER__` definition in the Linux kernel. That's all about useful meta-information related to our `sys_write` system call, let's try to understand how this system call is implemented. As we can see it starts from the definition of the `f` structure that has `fd` structure type that represents file descriptor in the Linux kernel and we put the result of the call of the `fdget_pos` function. The `fdget_pos` function defined in the same source code file and just expands the call of the `__to_fd` function:

```
static inline struct fd fdget_pos(int fd)
{
        return __to_fd(__fdget_pos(fd));
}
```

The main purpose of the `fdget_pos` is to convert the given file descriptor which is just a number to the `fd` structure. Through the long chain of function calls, the `fdget_pos` function gets the file descriptor table of the current process, `current->files` , and tries to find a corresponding file descriptor number there. As we got the `fd` structure for the given file descriptor number, we check it and return if it does not exist. We get the current position in the file with the call of the `file_pos_read` function that just returns `f_pos` field of our file:

```
static inline loff_t file_pos_read(struct file *file)
{
        return file->f_pos;
}
```

and calls the `vfs_write` function. The `vfs_write` function defined in the fs/read_write.c source code file and does the work for us - writes given buffer to the given file starting from the given position. We will not dive into details about the `vfs_write` function, because this function is weakly related to the `system call` concept but mostly about Virtual file system concept which we will see in another chapter. After the `vfs_write` has finished its work, we check the result and if it was finished successfully we change the position in the file with the `file_pos_write` function:

```
if (ret >= 0)
    file_pos_write(f.file, pos);
```

that just updates `f_pos` with the given position in the given file:

```
static inline void file_pos_write(struct file *file, loff_t pos)
{
        file->f_pos = pos;
}
```

At the end of the our `write` system call handler, we can see the call of the following function:

```
fdput_pos(f);
```

unlocks the `f_pos_lock` mutex that protects file position during concurrent writes from threads that share file descriptor.

That's all.

We have seen the partial implementation of one system call provided by the Linux kernel. Of course we have missed some parts in the implementation of the `write` system call, because as I mentioned above, we will see only system calls related stuff in this chapter and will not see other stuff related to other subsystems, such as Virtual file system.

Conclusion

This concludes the first part covering system call concepts in the Linux kernel. We have covered the theory of system calls so far and in the next part we will continue to dive into this topic, touching Linux kernel code related to system calls.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- system call
- vdso
- vsyscall
- general purpose registers
- socket
- C programming language
- x86
- x86_64
- x86-64 calling conventions
- System V Application Binary Interface. PDF
- GCC
- Intel manual. PDF
- system call table
- GCC macro documentation
- file descriptor
- stdout
- strace
- standard library
- wrapper functions
- ltrace
- sparse
- proc file system
- Virtual file system
- systemd
- epoll
- Previous chapter

System calls in the Linux kernel. Part 2.

How does the Linux kernel handle a system call

The previous part was the first part of the chapter that describes the system call concepts in the Linux kernel. In the previous part we learned what a system call is in the Linux kernel, and in operating systems in general. This was introduced from a user-space perspective, and part of the write system call implementation was discussed. In this part we continue our look at system calls, starting with some theory before moving onto the Linux kernel code.

A user application does not make the system call directly from our applications. We did not write the `Hello world!` program like:

```
int main(int argc, char **argv)
{
    ...
    ...
    ...
    sys_write(fd1, buf, strlen(buf));
    ...
    ...
}
```

We can use something similar with the help of C standard library and it will look something like this:

```
#include <unistd.h>

int main(int argc, char **argv)
{
    ...
    ...
    ...
    write(fd1, buf, strlen(buf));
    ...
    ...
}
```

But anyway, `write` is not a direct system call and not a kernel function. An application must fill general purpose registers with the correct values in the correct order and use the `syscall` instruction to make the actual system call. In this part we will look at what occurs in the Linux kernel when the `syscall` instruction is met by the processor.

Initialization of the system calls table

From the previous part we know that system call concept is very similar to an interrupt. Furthermore, system calls are implemented as software interrupts. So, when the processor handles a `syscall` instruction from a user application, this instruction causes an exception which transfers control to an exception handler. As we know, all exception handlers (or in other words kernel C functions that will react on an exception) are placed in the kernel code. But how does the Linux kernel search for the address of the necessary system call handler for the related system call? The Linux kernel contains a special table called the `system call table`. The system call table is represented by the `sys_call_table` array in the Linux kernel which is defined in the arch/x86/entry/syscall_64.c source code file. Let's look at its implementation:

```
asmlinkage const sys_call_ptr_t sys_call_table[__NR_syscall_max+1] = {
    [0 ... __NR_syscall_max] = &sys_ni_syscall,
    #include <asm/syscalls_64.h>
```

```
};
```

As we can see, the `sys_call_table` is an array of `__NR_syscall_max + 1` size where the `__NR_syscall_max` macro represents the maximum number of system calls for the given architecture. This book is about the x86_64 architecture, so for our case the `__NR_syscall_max` is `322` and this is the correct number at the time of writing (current Linux kernel version is `4.2.0-rc8+`). We can see this macro in the header file generated by Kbuild during kernel compilation - include/generated/asm-offsets.h`:

```
#define __NR_syscall_max 322
```

There will be the same number of system calls in the arch/x86/entry/syscalls/syscall_64.tbl for the `x86_64` . There are two important topics here; the type of the `sys_call_table` array, and the initialization of elements in this array. First of all, the type. The `sys_call_ptr_t` represents a pointer to a system call table. It is defined as typedef for a function pointer that returns nothing and does not take arguments:

```
typedef void (*sys_call_ptr_t)(void);
```

The second thing is the initialization of the `sys_call_table` array. As we can see in the code above, all elements of our array that contain pointers to the system call handlers point to the `sys_ni_syscall` . The `sys_ni_syscall` function represents not-implemented system calls. To start with, all elements of the `sys_call_table` array point to the not-implemented system call. This is the correct initial behaviour, because we only initialize storage of the pointers to the system call handlers, it is populated later on. Implementation of the `sys_ni_syscall` is pretty easy, it just returns -errno or `-ENOSYS` in our case:

```
asmlinkage long sys_ni_syscall(void)
{
    return -ENOSYS;
}
```

The `-ENOSYS` error tells us that:

```
ENOSYS          Function not implemented (POSIX.1)
```

Also a note on `...` in the initialization of the `sys_call_table` . We can do it with a GCC compiler extension called - Designated Initializers. This extension allows us to initialize elements in non-fixed order. As you can see, we include the `asm/syscalls_64.h` header at the end of the array. This header file is generated by the special script at arch/x86/entry/syscalls/syscalltbl.sh and generates our header file from the syscall table. The `asm/syscalls_64.h` contains definitions of the following macros:

```
__SYSCALL_COMMON(0, sys_read, sys_read)
__SYSCALL_COMMON(1, sys_write, sys_write)
__SYSCALL_COMMON(2, sys_open, sys_open)
__SYSCALL_COMMON(3, sys_close, sys_close)
__SYSCALL_COMMON(5, sys_newfstat, sys_newfstat)
...
...
...
```

The `__SYSCALL_COMMON` macro is defined in the same source code file and expands to the `__SYSCALL_64` macro which expands to the function definition:

```
#define __SYSCALL_COMMON(nr, sym, compat) __SYSCALL_64(nr, sym, compat)
#define __SYSCALL_64(nr, sym, compat) [nr] = sym,
```

So, after this, our `sys_call_table` takes the following form:

```
asmlinkage const sys_call_ptr_t sys_call_table[__NR_syscall_max+1] = {
    [0 ... __NR_syscall_max] = &sys_ni_syscall,
    [0] = sys_read,
    [1] = sys_write,
    [2] = sys_open,
    ...
    ...
    ...
};
```

After this all elements that point to the non-implemented system calls will contain the address of the `sys_ni_syscall` function that just returns `-ENOSYS` as we saw above, and other elements will point to the `sys_syscall_name` functions.

At this point, we have filled the system call table and the Linux kernel knows where each system call handler is. But the Linux kernel does not call a `sys_syscall_name` function immediately after it is instructed to handle a system call from a user space application. Remember the chapter about interrupts and interrupt handling. When the Linux kernel gets the control to handle an interrupt, it had to do some preparations like save user space registers, switch to a new stack and many more tasks before it will call an interrupt handler. There is the same situation with the system call handling. The preparation for handling a system call is the first thing, but before the Linux kernel will start these preparations, the entry point of a system call must be initialized and only the Linux kernel knows how to perform this preparation. In the next paragraph we will see the process of the initialization of the system call entry in the Linux kernel.

Initialization of the system call entry

When a system call occurs in the system, where are the first bytes of code that starts to handle it? As we can read in the Intel manual - 64-ia-32-architectures-software-developer-vol-2b-manual:

```
SYSCALL invokes an OS system-call handler at privilege level 0.
It does so by loading RIP from the IA32_LSTAR MSR
```

it means that we need to put the system call entry in to the `IA32_LSTAR` model specific register. This operation takes place during the Linux kernel initialization process. If you have read the fourth part of the chapter that describes interrupts and interrupt handling in the Linux kernel, you know that the Linux kernel calls the `trap_init` function during the initialization process. This function is defined in the arch/x86/kernel/setup.c source code file and executes the initialization of the `non-early` exception handlers like divide error, coprocessor error etc. Besides the initialization of the `non-early` exceptions handlers, this function calls the `cpu_init` function from the arch/x86/kernel/cpu/common.c source code file which besides initialization of `per-cpu` state, calls the `syscall_init` function from the same source code file.

This function performs the initialization of the system call entry point. Let's look on the implementation of this function. It does not take parameters and first of all it fills two model specific registers:

```
wrmsrl(MSR_STAR,  ((u64)__USER32_CS)<<48  | ((u64)__KERNEL_CS)<<32);
wrmsrl(MSR_LSTAR, entry_SYSCALL_64);
```

The first model specific register - `MSR_STAR` contains `63:48` bits of the user code segment. These bits will be loaded to the `CS` and `SS` segment registers for the `sysret` instruction which provides functionality to return from a system call to user code with the related privilege. Also the `MSR_STAR` contains `47:32` bits from the kernel code that will be used as the base selector for `CS` and `SS` segment registers when user space applications execute a system call. In the second line of code we fill the `MSR_LSTAR` register with the `entry_SYSCALL_64` symbol that represents system call entry. The `entry_SYSCALL_64` is defined in the arch/x86/entry/entry_64.S assembly file and contains code related to the preparation performed before a system call handler will be executed (I already wrote about these preparations, read above). We will not consider the `entry_SYSCALL_64` now, but will return to it later in this chapter.

After we have set the entry point for system calls, we need to set the following model specific registers:

- `MSR_CSTAR` - target `rip` for the compatibility mode callers;
- `MSR_IA32_SYSENTER_CS` - target `cs` for the `sysenter` instruction;
- `MSR_IA32_SYSENTER_ESP` - target `esp` for the `sysenter` instruction;
- `MSR_IA32_SYSENTER_EIP` - target `eip` for the `sysenter` instruction.

The values of these model specific register depend on the `CONFIG_IA32_EMULATION` kernel configuration option. If this kernel configuration option is enabled, it allows legacy 32-bit programs to run under a 64-bit kernel. In the first case, if the `CONFIG_IA32_EMULATION` kernel configuration option is enabled, we fill these model specific registers with the entry point for the system calls the compatibility mode:

```
wrmsrl(MSR_CSTAR, entry_SYSCALL_compat);
```

and with the kernel code segment, put zero to the stack pointer and write the address of the `entry_SYSENTER_compat` symbol to the instruction pointer:

```
wrmsrl_safe(MSR_IA32_SYSENTER_CS, (u64)__KERNEL_CS);
wrmsrl_safe(MSR_IA32_SYSENTER_ESP, 0ULL);
wrmsrl_safe(MSR_IA32_SYSENTER_EIP, (u64)entry_SYSENTER_compat);
```

In another way, if the `CONFIG_IA32_EMULATION` kernel configuration option is disabled, we write `ignore_sysret` symbol to the `MSR_CSTAR` :

```
wrmsrl(MSR_CSTAR, ignore_sysret);
```

that is defined in the arch/x86/entry/entry_64.S assembly file and just returns `-ENOSYS` error code:

```
ENTRY(ignore_sysret)
    mov    $-ENOSYS, %eax
    sysret
END(ignore_sysret)
```

Now we need to fill `MSR_IA32_SYSENTER_CS` , `MSR_IA32_SYSENTER_ESP` , `MSR_IA32_SYSENTER_EIP` model specific registers as we did in the previous code when the `CONFIG_IA32_EMULATION` kernel configuration option was enabled. In this case (when the `CONFIG_IA32_EMULATION` configuration option is not set) we fill the `MSR_IA32_SYSENTER_ESP` and the `MSR_IA32_SYSENTER_EIP` with zero and put the invalid segment of the Global Descriptor Table to the `MSR_IA32_SYSENTER_CS` model specific register:

```
wrmsrl_safe(MSR_IA32_SYSENTER_CS, (u64)GDT_ENTRY_INVALID_SEG);
wrmsrl_safe(MSR_IA32_SYSENTER_ESP, 0ULL);
wrmsrl_safe(MSR_IA32_SYSENTER_EIP, 0ULL);
```

You can read more about the `Global Descriptor Table` in the second part of the chapter that describes the booting process of the Linux kernel.

At the end of the `syscall_init` function, we just mask flags in the flags register by writing the set of flags to the `MSR_SYSCALL_MASK` model specific register:

```
wrmsrl(MSR_SYSCALL_MASK,
       X86_EFLAGS_TF|X86_EFLAGS_DF|X86_EFLAGS_IF|
       X86_EFLAGS_IOPL|X86_EFLAGS_AC|X86_EFLAGS_NT);
```

These flags will be cleared during syscall initialization. That's all, it is the end of the `syscall_init` function and it means that system call entry is ready to work. Now we can see what will occur when a user application executes the `syscall` instruction.

Preparation before system call handler will be called

As I already wrote, before a system call or an interrupt handler will be called by the Linux kernel we need to do some preparations. The `idtentry` macro performs the preparations required before an exception handler will be executed, the `interrupt` macro performs the preparations required before an interrupt handler will be called and the `entry_SYSCALL_64` will do the preparations required before a system call handler will be executed.

The `entry_SYSCALL_64` is defined in the arch/x86/entry/entry_64.S assembly file and starts from the following macro:

```
SWAPGS_UNSAFE_STACK
```

This macro is defined in the arch/x86/include/asm/irqflags.h header file and expands to the `swapgs` instruction:

```
#define SWAPGS_UNSAFE_STACK    swapgs
```

which exchanges the current GS base register value with the value contained in the `MSR_KERNEL_GS_BASE` model specific register. In other words we moved it on to the kernel stack. After this we point the old stack pointer to the `rsp_scratch` per-cpu variable and setup the stack pointer to point to the top of stack for the current processor:

```
movq    %rsp, PER_CPU_VAR(rsp_scratch)
movq    PER_CPU_VAR(cpu_current_top_of_stack), %rsp
```

In the next step we push the stack segment and the old stack pointer to the stack:

```
pushq    $__USER_DS
pushq    PER_CPU_VAR(rsp_scratch)
```

After this we enable interrupts, because interrupts are `off` on entry and save the general purpose registers (besides `bp` , `bx` and from `r12` to `r15`), flags, `-ENOSYS` for the non-implemented system call and code segment register on the stack:

```
ENABLE_INTERRUPTS(CLBR_NONE)

pushq    %r11
pushq    $__USER_CS
pushq    %rcx
pushq    %rax
pushq    %rdi
pushq    %rsi
pushq    %rdx
pushq    %rcx
pushq    $-ENOSYS
pushq    %r8
pushq    %r9
pushq    %r10
pushq    %r11
sub    $(6*8), %rsp
```

When a system call occurs from the user's application, general purpose registers have the following state:

- `rax` - contains system call number;
- `rcx` - contains return address to the user space;
- `r11` - contains register flags;
- `rdi` - contains first argument of a system call handler;
- `rsi` - contains second argument of a system call handler;
- `rdx` - contains third argument of a system call handler;

* `r10` - contains fourth argument of a system call handler;
* `r8` - contains fifth argument of a system call handler;
* `r9` - contains sixth argument of a system call handler;

Other general purpose registers (as `rbp` , `rbx` and from `r12` to `r15`) are callee-preserved in C ABI). So we push register flags on the top of the stack, then user code segment, return address to the user space, system call number, first three arguments, dump error code for the non-implemented system call and other arguments on the stack.

In the next step we check the `_TIF_WORK_SYSCALL_ENTRY` in the current `thread_info` :

```
testl    $_TIF_WORK_SYSCALL_ENTRY, ASM_THREAD_INFO(TI_flags, %rsp, SIZEOF_PTREGS)
jnz    tracesys
```

The `_TIF_WORK_SYSCALL_ENTRY` macro is defined in the arch/x86/include/asm/thread_info.h header file and provides set of the thread information flags that are related to the system calls tracing:

```
#define _TIF_WORK_SYSCALL_ENTRY \
    (_TIF_SYSCALL_TRACE | _TIF_SYSCALL_EMU | _TIF_SYSCALL_AUDIT |   \
    _TIF_SECCOMP | _TIF_SINGLESTEP | _TIF_SYSCALL_TRACEPOINT |     \
    _TIF_NOHZ)
```

We will not consider debugging/tracing related stuff in this chapter, but will see it in the separate chapter that will be devoted to the debugging and tracing techniques in the Linux kernel. After the `tracesys` label, the next label is the `entry_SYSCALL_64_fastpath` . In the `entry_SYSCALL_64_fastpath` we check the `__SYSCALL_MASK` that is defined in the arch/x86/include/asm/unistd.h header file and

```
# ifdef CONFIG_X86_X32_ABI
#  define __SYSCALL_MASK (~(__X32_SYSCALL_BIT))
# else
#  define __SYSCALL_MASK (~0)
# endif
```

where the `__X32_SYSCALL_BIT` is

```
#define __X32_SYSCALL_BIT    0x40000000
```

As we can see the `__SYSCALL_MASK` depends on the `CONFIG_X86_X32_ABI` kernel configuration option and represents the mask for the 32-bit ABI in the 64-bit kernel.

So we check the value of the `__SYSCALL_MASK` and if the `CONFIG_X86_X32_ABI` is disabled we compare the value of the `rax` register to the maximum syscall number (`__NR_syscall_max`), alternatively if the `CONFIG_X86_X32_ABI` is enabled we mask the `eax` register with the `__X32_SYSCALL_BIT` and do the same comparison:

```
#if __SYSCALL_MASK == ~0
    cmpq    $__NR_syscall_max, %rax
#else
    andl    $__SYSCALL_MASK, %eax
    cmpl    $__NR_syscall_max, %eax
#endif
```

After this we check the result of the last comparison with the `ja` instruction that executes if `CF` and `ZF` flags are zero:

```
ja    1f
```

and if we have the correct system call for this, we move the fourth argument from the `r10` to the `rcx` to keep x86_64 C ABI compliant and execute the `call` instruction with the address of a system call handler:

```
movq    %r10, %rcx
call    *sys_call_table(, %rax, 8)
```

Note, the `sys_call_table` is an array that we saw above in this part. As we already know the `rax` general purpose register contains the number of a system call and each element of the `sys_call_table` is 8-bytes. So we are using `*sys_call_table(, %rax, 8)` this notation to find the correct offset in the `sys_call_table` array for the given system call handler.

That's all. We did all the required preparations and the system call handler was called for the given interrupt handler, for example `sys_read`, `sys_write` or other system call handler that is defined with the `SYSCALL_DEFINE[N]` macro in the Linux kernel code.

Exit from a system call

After a system call handler finishes its work, we will return back to the arch/x86/entry/entry_64.S, right after where we have called the system call handler:

```
call    *sys_call_table(, %rax, 8)
```

The next step after we've returned from a system call handler is to put the return value of a system handler on to the stack. We know that a system call returns the result to the user program in the general purpose `rax` register, so we are moving its value on to the stack after the system call handler has finished its work:

```
movq    %rax, RAX(%rsp)
```

on the `RAX` place.

After this we can see the call of the `LOCKDEP_SYS_EXIT` macro from the arch/x86/include/asm/irqflags.h:

```
LOCKDEP_SYS_EXIT
```

The implementation of this macro depends on the `CONFIG_DEBUG_LOCK_ALLOC` kernel configuration option that allows us to debug locks on exit from a system call. And again, we will not consider it in this chapter, but will return to it in a separate one. In the end of the `entry_SYSCALL_64` function we restore all general purpose registers besides `rcx` and `r11`, because the `rcx` register must contain the return address to the application that called system call and the `r11` register contains the old flags register. After all general purpose registers are restored, we fill `rcx` with the return address, `r11` register with the flags and `rsp` with the old stack pointer:

```
RESTORE_C_REGS_EXCEPT_RCX_R11

movq    RIP(%rsp), %rcx
movq    EFLAGS(%rsp), %r11
movq    RSP(%rsp), %rsp

USERGS_SYSRET64
```

In the end we just call the `USERGS_SYSRET64` macro that expands to the call of the `swapgs` instruction which exchanges again the user `GS` and kernel `GS` and the `sysretq` instruction which executes on exit from a system call handler:

```
#define USERGS_SYSRET64                \
    swapgs;                            \
```

```
    sysretq;
```

Now we know what occurs when a user application calls a system call. The full path of this process is as follows:

- User application contains code that fills general purpose register with the values (system call number and arguments of this system call);
- Processor switches from the user mode to kernel mode and starts execution of the system call entry - `entry_SYSCALL_64` ;
- `entry_SYSCALL_64` switches to the kernel stack and saves some general purpose registers, old stack and code segment, flags and etc... on the stack;
- `entry_SYSCALL_64` checks the system call number in the `rax` register, searches a system call handler in the `sys_call_table` and calls it, if the number of a system call is correct;
- If a system call is not correct, jump on exit from system call;
- After a system call handler will finish its work, restore general purpose registers, old stack, flags and return address and exit from the `entry_SYSCALL_64` with the `sysretq` instruction.

That's all.

Conclusion

This is the end of the second part about the system calls concept in the Linux kernel. In the previous part we saw theory about this concept from the user application view. In this part we continued to dive into the stuff which is related to the system call concept and saw what the Linux kernel does when a system call occurs.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- system call
- write
- C standard library
- list of cpu architectures
- x86_64
- kbuild
- typedef
- errno
- gcc
- model specific register
- intel 2b manual
- coprocessor
- instruction pointer
- flags register
- Global Descriptor Table
- per-cpu
- general purpose registers
- ABI
- x86_64 C ABI
- previous chapter

System calls in the Linux kernel. Part 3.

vsyscalls and vDSO

This is the third part of the chapter that describes system calls in the Linux kernel and we saw preparations after a system call caused by a userspace application and process of handling of a system call in the previous part. In this part we will look at two concepts that are very close to the system call concept, they are called `vsyscall` and `vdso` .

We already know what `system call` s are. They are special routines in the Linux kernel which userspace applications ask to do privileged tasks, like to read or to write to a file, to open a socket, etc. As you may know, invoking a system call is an expensive operation in the Linux kernel, because the processor must interrupt the currently executing task and switch context to kernel mode, subsequently jumping again into userspace after the system call handler finishes its work. These two mechanisms - `vsyscall` and `vdso` are designed to speed up this process for certain system calls and in this part we will try to understand how these mechanisms work.

Introduction to vsyscalls

The `vsyscall` or `virtual system call` is the first and oldest mechanism in the Linux kernel that is designed to accelerate execution of certain system calls. The principle of work of the `vsyscall` concept is simple. The Linux kernel maps into user space a page that contains some variables and the implementation of some system calls. We can find information about this memory space in the Linux kernel documentation for the x86_64:

```
ffffffffff600000 - ffffffffffdfffff (=8 MB) vsyscalls
```

or:

```
~$ sudo cat /proc/1/maps | grep vsyscall
ffffffffff600000-ffffffffff601000 r-xp 00000000 00:00 0                  [vsyscall]
```

After this, these system calls will be executed in userspace and this means that there will not be context switching. Mapping of the `vsyscall` page occurs in the `map_vsyscall` function that is defined in the arch/x86/entry/vsyscall/vsyscall_64.c source code file. This function is called during the Linux kernel initialization in the `setup_arch` function that is defined in the arch/x86/kernel/setup.c source code file (we saw this function in the fifth part of the Linux kernel initialization process chapter).

Note that implementation of the `map_vsyscall` function depends on the `CONFIG_X86_VSYSCALL_EMULATION` kernel configuration option:

```
#ifdef CONFIG_X86_VSYSCALL_EMULATION
extern void map_vsyscall(void);
#else
static inline void map_vsyscall(void) {}
#endif
```

As we can read in the help text, the `CONFIG_X86_VSYSCALL_EMULATION` configuration option: `Enable vsyscall emulation` . Why emulate `vsyscall` ? Actually, the `vsyscall` is a legacy ABI due to security reasons. Virtual system calls have fixed addresses, meaning that `vsyscall` page is still at the same location every time and the location of this page is determined in the `map_vsyscall` function. Let's look on the implementation of this function:

```
void __init map_vsyscall(void)
{
```

```
    extern char __vsyscall_page;
    unsigned long physaddr_vsyscall = __pa_symbol(&__vsyscall_page);
    ...
    ...
    ...
}
```

As we can see, at the beginning of the `map_vsyscall` function we get the physical address of the `vsyscall` page with the `__pa_symbol` macro (we already saw implementation if this macro in the fourth path of the Linux kernel initialization process). The `__vsyscall_page` symbol defined in the arch/x86/entry/vsyscall/vsyscall_emu_64.S assembly source code file and have the following virtual address:

```
ffffffff81881000 D __vsyscall_page
```

in the `.data..page_aligned, aw` section and contains call of the three following system calls:

- `gettimeofday` ;
- `time` ;
- `getcpu` .

Or:

```
__vsyscall_page:
    mov $__NR_gettimeofday, %rax
    syscall
    ret

    .balign 1024, 0xcc
    mov $__NR_time, %rax
    syscall
    ret

    .balign 1024, 0xcc
    mov $__NR_getcpu, %rax
    syscall
    ret
```

Let's go back to the implementation of the `map_vsyscall` function and return to the implementation of the `__vsyscall_page` later. After we received the physical address of the `__vsyscall_page` , we check the value of the `vsyscall_mode` variable and set the fix-mapped address for the `vsyscall` page with the `__set_fixmap` macro:

```
if (vsyscall_mode != NONE)
    __set_fixmap(VSYSCALL_PAGE, physaddr_vsyscall,
                 vsyscall_mode == NATIVE
                             ? PAGE_KERNEL_VSYSCALL
                             : PAGE_KERNEL_VVAR);
```

The `__set_fixmap` takes three arguments: The first is index of the `fixed_addresses` enum. In our case `VSYSCALL_PAGE` is the first element of the `fixed_addresses` enum for the `x86_64` architecture:

```
enum fixed_addresses {
...
...
...
#ifdef CONFIG_X86_VSYSCALL_EMULATION
    VSYSCALL_PAGE = (FIXADDR_TOP - VSYSCALL_ADDR) >> PAGE_SHIFT,
#endif
...
...
...
```

It equal to the `511`. The second argument is the physical address of the page that has to be mapped and the third argument is the flags of the page. Note that the flags of the `VSYSCALL_PAGE` depend on the `vsyscall_mode` variable. It will be `PAGE_KERNEL_VSYSCALL` if the `vsyscall_mode` variable is `NATIVE` and the `PAGE_KERNEL_VVAR` otherwise. Both macros (the `PAGE_KERNEL_VSYSCALL` and the `PAGE_KERNEL_VVAR`) will be expanded to the following flags:

```
#define __PAGE_KERNEL_VSYSCALL          (__PAGE_KERNEL_RX | _PAGE_USER)
#define __PAGE_KERNEL_VVAR              (__PAGE_KERNEL_RO | _PAGE_USER)
```

that represent access rights to the `vsyscall` page. Both flags have the same `_PAGE_USER` flags that means that the page can be accessed by a user-mode process running at lower privilege levels. The second flag depends on the value of the `vsyscall_mode` variable. The first flag (`__PAGE_KERNEL_VSYSCALL`) will be set in the case where `vsyscall_mode` is `NATIVE`. This means virtual system calls will be native `syscall` instructions. In other way the vsyscall will have `PAGE_KERNEL_VVAR` if the `vsyscall_mode` variable will be `emulate`. In this case virtual system calls will be turned into traps and are emulated reasonably. The `vsyscall_mode` variable gets its value in the `vsyscall_setup` function:

```
static int __init vsyscall_setup(char *str)
{
    if (str) {
        if (!strcmp("emulate", str))
            vsyscall_mode = EMULATE;
        else if (!strcmp("native", str))
            vsyscall_mode = NATIVE;
        else if (!strcmp("none", str))
            vsyscall_mode = NONE;
        else
            return -EINVAL;

        return 0;
    }

    return -EINVAL;
}
```

That will be called during early kernel parameters parsing:

```
early_param("vsyscall", vsyscall_setup);
```

More about `early_param` macro you can read in the sixth part of the chapter that describes process of the initialization of the Linux kernel.

In the end of the `vsyscall_map` function we just check that virtual address of the `vsyscall` page is equal to the value of the `VSYSCALL_ADDR` with the BUILD_BUG_ON macro:

```
BUILD_BUG_ON((unsigned long)__fix_to_virt(VSYSCALL_PAGE) !=
             (unsigned long)VSYSCALL_ADDR);
```

That's all. `vsyscall` page is set up. The result of the all the above is the following: If we pass `vsyscall=native` parameter to the kernel command line, virtual system calls will be handled as native `syscall` instructions in the arch/x86/entry/vsyscall/vsyscall_emu_64.S. The glibc knows addresses of the virtual system call handlers. Note that virtual system call handlers are aligned by `1024` (or `0x400`) bytes:

```
__vsyscall_page:
    mov $__NR_gettimeofday, %rax
    syscall
    ret
```

```
    .balign 1024, 0xcc
    mov $__NR_time, %rax
    syscall
    ret

    .balign 1024, 0xcc
    mov $__NR_getcpu, %rax
    syscall
    ret
```

And the start address of the `vsyscall` page is the `ffffffffff600000` every time. So, the glibc knows the addresses of the all virtual system call handlers. You can find definition of these addresses in the `glibc` source code:

```
#define VSYSCALL_ADDR_vgettimeofday   0xffffffffff600000
#define VSYSCALL_ADDR_vtime           0xffffffffff600400
#define VSYSCALL_ADDR_vgetcpu          0xffffffffff600800
```

All virtual system call requests will fall into the `__vsyscall_page` + `VSYSCALL_ADDR_vsyscall_name` offset, put the number of a virtual system call to the `rax` general purpose register and the native for the x86_64 `syscall` instruction will be executed.

In the second case, if we pass `vsyscall=emulate` parameter to the kernel command line, an attempt to perform virtual system call handler will cause a page fault exception. Of course, remember, the `vsyscall` page has `__PAGE_KERNEL_VVAR` access rights that forbid execution. The `do_page_fault` function is the `#PF` or page fault handler. It tries to understand the reason of the last page fault. And one of the reason can be situation when virtual system call called and `vsyscall` mode is `emulate`. In this case `vsyscall` will be handled by the `emulate_vsyscall` function that defined in the arch/x86/entry/vsyscall/vsyscall_64.c source code file.

The `emulate_vsyscall` function gets the number of a virtual system call, checks it, prints error and sends segmentation fault simply:

```
...
...
...
vsyscall_nr = addr_to_vsyscall_nr(address);
if (vsyscall_nr < 0) {
    warn_bad_vsyscall(KERN_WARNING, regs, "misaligned vsyscall...);
    goto sigsegv;
}
...
...
...
sigsegv:
    force_sig(SIGSEGV, current);
    reutrn true;
```

As it checked number of a virtual system call, it does some yet another checks like `access_ok` violations and execute system call function depends on the number of a virtual system call:

```
switch (vsyscall_nr) {
    case 0:
        ret = sys_gettimeofday(
            (struct timeval __user *)regs->di,
            (struct timezone __user *)regs->si);
        break;
    ...
    ...
    ...
}
```

In the end we put the result of the `sys_gettimeofday` or another virtual system call handler to the `ax` general purpose register, as we did it with the normal system calls and restore the instruction pointer register and add `8` bytes to the stack pointer register. This operation emulates `ret` instruction.

```
    regs->ax = ret;

do_ret:
    regs->ip = caller;
    regs->sp += 8;
    return true;
```

That's all. Now let's look on the modern concept - `vDSO` .

Introduction to vDSO

As I already wrote above, `vsyscall` is an obsolete concept and replaced by the `vDSO` or `virtual dynamic shared object` . The main difference between the `vsyscall` and `vDSO` mechanisms is that `vDSO` maps memory pages into each process in a shared object form, but `vsyscall` is static in memory and has the same address every time. For the `x86_64` architecture it is called - `linux-vdso.so.1` . All userspace applications that dynamically link to `glibc` will use the `vDSO` automatically. For example:

```
~$ ldd /bin/uname
    linux-vdso.so.1 (0x00007ffe014b7000)
    libc.so.6 => /lib64/libc.so.6 (0x00007fbfee2fe000)
    /lib64/ld-linux-x86-64.so.2 (0x00005559aab7c000)
```

Or:

```
~$ sudo cat /proc/1/maps | grep vdso
7fff39f73000-7fff39f75000 r-xp 00000000 00:00 0       [vdso]
```

Here we can see that uname util was linked with the three libraries:

- `linux-vdso.so.1` ;
- `libc.so.6` ;
- `ld-linux-x86-64.so.2` .

The first provides `vDSO` functionality, the second is `C` standard library and the third is the program interpreter (more about this you can read in the part that describes linkers). So, the `vDSO` solves limitations of the `vsyscall` . Implementation of the `vDSO` is similar to `vsyscall` .

Initialization of the `vDSO` occurs in the `init_vdso` function that defined in the arch/x86/entry/vdso/vma.c source code file. This function starts from the initialization of the `vDSO` images for 32-bits and 64-bits depends on the `CONFIG_X86_X32_ABI` kernel configuration option:

```
static int __init init_vdso(void)
{
    init_vdso_image(&vdso_image_64);

#ifdef CONFIG_X86_X32_ABI
    init_vdso_image(&vdso_image_x32);
#endif
```

Both functions initialize the `vdso_image` structure. This structure is defined in the two generated source code files: the arch/x86/entry/vdso/vdso-image-64.c and the arch/x86/entry/vdso/vdso-image-64.c. These source code files generated by the vdso2c program from the different source code files, represent different approaches to call a system call like `int 0x80`, `sysenter` and etc. The full set of the images depends on the kernel configuration.

For example for the `x86_64` Linux kernel it will contain `vdso_image_64`:

```
#ifdef CONFIG_X86_64
extern const struct vdso_image vdso_image_64;
#endif
```

But for the `x86` - `vdso_image_32`:

```
#ifdef CONFIG_X86_X32
extern const struct vdso_image vdso_image_x32;
#endif
```

If our kernel is configured for the `x86` architecture or for the `x86_64` and compatibility mode, we will have ability to call a system call with the `int 0x80` interrupt, if compatibility mode is enabled, we will be able to call a system call with the native `syscall instruction` or `sysenter` instruction in other way:

```
#if defined CONFIG_X86_32 || defined CONFIG_COMPAT
  extern const struct vdso_image vdso_image_32_int80;
#ifdef CONFIG_COMPAT
  extern const struct vdso_image vdso_image_32_syscall;
#endif
 extern const struct vdso_image vdso_image_32_sysenter;
#endif
```

As we can understand from the name of the `vdso_image` structure, it represents image of the `vDSO` for the certain mode of the system call entry. This structure contains information about size in bytes of the `vDSO` area that always a multiple of `PAGE_SIZE` (`4096` bytes), pointer to the text mapping, start and end address of the `alternatives` (set of instructions with better alternatives for the certain type of the processor) and etc. For example `vdso_image_64` looks like this:

```
const struct vdso_image vdso_image_64 = {
    .data = raw_data,
    .size = 8192,
    .text_mapping = {
        .name = "[vdso]",
        .pages = pages,
    },
    .alt = 3145,
    .alt_len = 26,
    .sym_vvar_start = -8192,
    .sym_vvar_page = -8192,
    .sym_hpet_page = -4096,
};
```

Where the `raw_data` contains raw binary code of the 64-bit `vDSO` system calls which are `2` page size:

```
static struct page *pages[2];
```

or 8 Kilobytes.

The `init_vdso_image` function is defined in the same source code file and just initializes the `vdso_image.text_mapping.pages`. First of all this function calculates the number of pages and initializes each `vdso_image.text_mapping.pages[number_of_page]` with the `virt_to_page` macro that converts given address to the `page` structure:

```
void __init init_vdso_image(const struct vdso_image *image)
{
    int i;
    int npages = (image->size) / PAGE_SIZE;

    for (i = 0; i < npages; i++)
        image->text_mapping.pages[i] =
            virt_to_page(image->data + i*PAGE_SIZE);
    ...
    ...
    ...
}
```

The `init_vdso` function passed to the `subsys_initcall` macro adds the given function to the `initcalls` list. All functions from this list will be called in the `do_initcalls` function from the init/main.c source code file:

```
subsys_initcall(init_vdso);
```

Ok, we just saw initialization of the `vDSO` and initialization of `page` structures that are related to the memory pages that contain `vDSO` system calls. But to where do their pages map? Actually they are mapped by the kernel, when it loads binary to the memory. The Linux kernel calls the `arch_setup_additional_pages` function from the arch/x86/entry/vdso/vma.c source code file that checks that `vDSO` enabled for the `x86_64` and calls the `map_vdso` function:

```
int arch_setup_additional_pages(struct linux_binprm *bprm, int uses_interp)
{
    if (!vdso64_enabled)
        return 0;

    return map_vdso(&vdso_image_64, true);
}
```

The `map_vdso` function is defined in the same source code file and maps pages for the `vDSO` and for the shared `vDSO` variables. That's all. The main differences between the `vsyscall` and the `vDSO` concepts is that `vsyscall` has a static address of `ffffffffff600000` and implements `3` system calls, whereas the `vDSO` loads dynamically and implements four system calls:

- `__vdso_clock_gettime` ;
- `__vdso_getcpu` ;
- `__vdso_gettimeofday` ;
- `__vdso_time` .

That's all.

Conclusion

This is the end of the third part about the system calls concept in the Linux kernel. In the previous part we discussed the implementation of the preparation from the Linux kernel side, before a system call will be handled and implementation of the `exit` process from a system call handler. In this part we continued to dive into the stuff which is related to the system call concept and learned two new concepts that are very similar to the system call - the `vsyscall` and the `vDSO` .

After all of these three parts, we know almost all things that are related to system calls, we know what system call is and why user applications need them. We also know what occurs when a user application calls a system call and how the kernel handles system calls.

The next part will be the last part in this chapter and we will see what occurs when a user runs the program.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- x86_64 memory map
- x86_64
- context switching
- ABI
- virtual address
- Segmentation
- enum
- fix-mapped addresses
- glibc
- BUILD_BUG_ON
- Processor register
- Page fault
- segmentation fault
- instruction pointer
- stack pointer
- uname
- Linkers
- Previous part

System calls in the Linux kernel. Part 4.

How does the Linux kernel run a program

This is the fourth part of the chapter that describes system calls in the Linux kernel and as I wrote in the conclusion of the previous - this part will be last in this chapter. In the previous part we stopped at the two new concepts:

- `vsyscall` ;
- `vDSO` ;

that are related and very similar on system call concept.

This part will be last part in this chapter and as you can understand from the part's title - we will see what does occur in the Linux kernel when we run our programs. So, let's start.

how do we launch our programs?

There are many different ways to launch an application from a user perspective. For example we can run a program from the shell or double-click on the application icon. It does not matter. The Linux kernel handles application launch regardless how we do launch this application.

In this part we will consider the way when we just launch an application from the shell. As you know, the standard way to launch an application from shell is the following: We just launch a terminal emulator application and just write the name of the program and pass or not arguments to our program, for example:

```
~$ ls --version
ls (GNU coreutils) 8.23
Copyright (C) 2014 Free Software Foundation, Inc.
License GPLv3+: GNU GPL version 3 or later <http://gnu.org/licenses/gpl.html>.
This is free software: you are free to change and redistribute it.
There is NO WARRANTY, to the extent permitted by law.

Written by Richard M. Stallman and David MacKenzie.
```

Let's consider what does occur when we launch an application from the shell, what does shell do when we write program name, what does Linux kernel do etc. But before we will start to consider these interesting things, I want to warn that this book is about the Linux kernel. That's why we will see Linux kernel insides related stuff mostly in this part. We will not consider in details what does shell do, we will not consider complex cases, for example subshells etc.

My default shell is - bash, so I will consider how do bash shell launches a program. So let's start. The `bash` shell as well as any program that written with C programming language starts from the main function. If you will look on the source code of the `bash` shell, you will find the `main` function in the shell.c source code file. This function makes many different things before the main thread loop of the `bash` started to work. For example this function:

- checks and tries to open `/dev/tty` ;
- check that shell running in debug mode;
- parses command line arguments;
- reads shell environment;
- loads `.bashrc` , `.profile` and other configuration files;
- and many many more.

After all of these operations we can see the call of the `reader_loop` function. This function defined in the eval.c source code file and represents main thread loop or in other words it reads and executes commands. As the `reader_loop` function made all checks and read the given program name and arguments, it calls the `execute_command` function from the execute_cmd.c source code file. The `execute_command` function through the chain of the functions calls:

```
execute_command
--> execute_command_internal
----> execute_simple_command
------> execute_disk_command
--------> shell_execve
```

makes different checks like do we need to start `subshell`, was it builtin `bash` function or not etc. As I already wrote above, we will not consider all details about things that are not related to the Linux kernel. In the end of this process, the `shell_execve` function calls the `execve` system call:

```
execve (command, args, env);
```

The `execve` system call has the following signature:

```
int execve(const char *filename, char *const argv [], char *const envp[]);
```

and executes a program by the given filename, with the given arguments and environment variables. This system call is the first in our case and only, for example:

```
$ strace ls
execve("/bin/ls", ["ls"], [/* 62 vars */]) = 0

$ strace echo
execve("/bin/echo", ["echo"], [/* 62 vars */]) = 0

$ strace uname
execve("/bin/uname", ["uname"], [/* 62 vars */]) = 0
```

So, a user application (`bash` in our case) calls the system call and as we already know the next step is Linux kernel.

execve system call

We saw preparation before a system call called by a user application and after a system call handler finished its work in the second part of this chapter. We stopped at the call of the `execve` system call in the previous paragraph. This system call defined in the fs/exec.c source code file and as we already know it takes three arguments:

```
SYSCALL_DEFINE3(execve,
        const char __user *, filename,
        const char __user *const __user *, argv,
        const char __user *const __user *, envp)
{
    return do_execve(getname(filename), argv, envp);
}
```

Implementation of the `execve` is pretty simple here, as we can see it just returns the result of the `do_execve` function. The `do_execve` function defined in the same source code file and do the following things:

- Initialize two pointers on a userspace data with the given arguments and environment variables;
- return the result of the `do_execveat_common` .

We can see its implementation:

```
struct user_arg_ptr argv = { .ptr.native = __argv };
struct user_arg_ptr envp = { .ptr.native = __envp };
return do_execveat_common(AT_FDCWD, filename, argv, envp, 0);
```

The `do_execveat_common` function does main work - it executes a new program. This function takes similar set of arguments, but as you can see it takes five arguments instead of three. The first argument is the file descriptor that represent directory with our application, in our case the `AT_FDCWD` means that the given pathname is interpreted relative to the current working directory of the calling process. The fifth argument is flags. In our case we passed `0` to the `do_execveat_common`. We will check in a next step, so will see it latter.

First of all the `do_execveat_common` function checks the `filename` pointer and returns if it is `NULL`. After this we check flags of the current process that limit of running processes is not exceed:

```
if (IS_ERR(filename))
    return PTR_ERR(filename);

if ((current->flags & PF_NPROC_EXCEEDED) &&
    atomic_read(&current_user()->processes) > rlimit(RLIMIT_NPROC)) {
    retval = -EAGAIN;
    goto out_ret;
}

current->flags &= ~PF_NPROC_EXCEEDED;
```

If these two checks were successful we unset `PF_NPROC_EXCEEDED` flag in the flags of the current process to prevent fail of the `execve`. You can see that in the next step we call the `unshare_files` function that defined in the kernel/fork.c and unshares the files of the current task and check the result of this function:

```
retval = unshare_files(&displaced);
if (retval)
    goto out_ret;
```

We need to call this function to eliminate potential leak of the execve'd binary's file descriptor. In the next step we start preparation of the `bprm` that represented by the `struct linux_binprm` structure (defined in the include/linux/binfmts.h header file). The `linux_binprm` structure is used to hold the arguments that are used when loading binaries. For example it contains `vma` field which has `vm_area_struct` type and represents single memory area over a contiguous interval in a given address space where our application will be loaded, `mm` field which is memory descriptor of the binary, pointer to the top of memory and many other different fields.

First of all we allocate memory for this structure with the `kzalloc` function and check the result of the allocation:

```
bprm = kzalloc(sizeof(*bprm), GFP_KERNEL);
if (!bprm)
    goto out_files;
```

After this we start to prepare the `binprm` credentials with the call of the `prepare_bprm_creds` function:

```
retval = prepare_bprm_creds(bprm);
    if (retval)
        goto out_free;

check_unsafe_exec(bprm);
current->in_execve = 1;
```

Initialization of the `binprm` credentials in other words is initialization of the `cred` structure that stored inside of the `linux_binprm` structure. The `cred` structure contains the security context of a task for example real uid of the task, real guid of the task, `uid` and `guid` for the virtual file system operations etc. In the next step as we executed preparation of the `bprm` credentials we check that now we can safely execute a program with the call of the `check_unsafe_exec` function and set the current process to the `in_execve` state.

After all of these operations we call the `do_open_execat` function that checks the flags that we passed to the `do_execveat_common` function (remember that we have `0` in the `flags`) and searches and opens executable file on disk, checks that our we will load a binary file from `noexec` mount points (we need to avoid execute a binary from filesystems that do not contain executable binaries like proc or sysfs), initializes `file` structure and returns pointer on this structure. Next we can see the call the `sched_exec` after this:

```
file = do_open_execat(fd, filename, flags);
retval = PTR_ERR(file);
if (IS_ERR(file))
    goto out_unmark;

sched_exec();
```

The `sched_exec` function is used to determine the least loaded processor that can execute the new program and to migrate the current process to it.

After this we need to check file descriptor of the give executable binary. We try to check does the name of the our binary file starts from the `/` symbol or does the path of the given executable binary is interpreted relative to the current working directory of the calling process or in other words file descriptor is `AT_FDCWD` (read above about this).

If one of these checks is successful we set the binary parameter filename:

```
bprm->file = file;

if (fd == AT_FDCWD || filename->name[0] == '/') {
    bprm->filename = filename->name;
}
```

Otherwise if the filename is empty we set the binary parameter filename to the `/dev/fd/%d` or `/dev/fd/%d/%s` depends on the filename of the given executable binary which means that we will execute the file to which the file descriptor refers:

```
} else {
    if (filename->name[0] == '\0')
        pathbuf = kasprintf(GFP_TEMPORARY, "/dev/fd/%d", fd);
    else
        pathbuf = kasprintf(GFP_TEMPORARY, "/dev/fd/%d/%s",
                            fd, filename->name);
    if (!pathbuf) {
        retval = -ENOMEM;
        goto out_unmark;
    }

    bprm->filename = pathbuf;
}

bprm->interp = bprm->filename;
```

Note that we set not only the `bprm->filename` but also `bprm->interp` that will contain name of the program interpreter. For now we just write the same name there, but later it will be updated with the real name of the program interpreter depends on binary format of a program. You can read above that we already prepared `cred` for the `linux_binprm` . The next step is initialization of other fields of the `linux_binprm` . First of all we call the `bprm_mm_init` function and pass the `bprm` to it:

```
retval = bprm_mm_init(bprm);
if (retval)
    goto out_unmark;
```

The `bprm_mm_init` defined in the same source code file and as we can understand from the function's name, it makes initialization of the memory descriptor or in other words the `bprm_mm_init` function initializes `mm_struct` structure. This structure defined in the include/linux/mm_types.h header file and represents address space of a process. We will not consider implementation of the `bprm_mm_init` function because we do not know many important stuff related to the Linux kernel memory manager, but we just need to know that this function initializes `mm_struct` and populate it with a temporary stack `vm_area_struct`.

After this we calculate the count of the command line arguments which are were passed to the our executable binary, the count of the environment variables and set it to the `bprm->argc` and `bprm->envc` respectively:

```
bprm->argc = count(argv, MAX_ARG_STRINGS);
if ((retval = bprm->argc) < 0)
    goto out;

bprm->envc = count(envp, MAX_ARG_STRINGS);
if ((retval = bprm->envc) < 0)
    goto out;
```

As you can see we do this operations with the help of the `count` function that defined in the same source code file and calculates the count of strings in the `argv` array. The `MAX_ARG_STRINGS` macro defined in the include/uapi/linux/binfmts.h header file and as we can understand from the macro's name, it represents maximum number of strings that were passed to the `execve` system call. The value of the `MAX_ARG_STRINGS`:

```
#define MAX_ARG_STRINGS 0x7FFFFFFF
```

After we calculated the number of the command line arguments and environment variables, we call the `prepare_binprm` function. We already call the function with the similar name before this moment. This function is called `prepare_binprm_cred` and we remember that this function initializes `cred` structure in the `linux_bprm`. Now the `prepare_binprm` function:

```
retval = prepare_binprm(bprm);
if (retval < 0)
    goto out;
```

fills the `linux_binprm` structure with the `uid` from inode and read `128` bytes from the binary executable file. We read only first `128` from the executable file because we need to check a type of our executable. We will read the rest of the executable file in the later step. After the preparation of the `linux_bprm` structure we copy the filename of the executable binary file, command line arguments and environment variables to the `linux_bprm` with the call of the `copy_strings_kernel` function:

```
retval = copy_strings_kernel(1, &bprm->filename, bprm);
if (retval < 0)
    goto out;

retval = copy_strings(bprm->envc, envp, bprm);
if (retval < 0)
    goto out;

retval = copy_strings(bprm->argc, argv, bprm);
if (retval < 0)
    goto out;
```

And set the pointer to the top of new program's stack that we set in the `bprm_mm_init` function:

```
bprm->exec = bprm->p;
```

The top of the stack will contain the program filename and we store this filename to the `exec` field of the `linux_bprm` structure.

Now we have filled `linux_bprm` structure, we call the `exec_binprm` function:

```
retval = exec_binprm(bprm);
if (retval < 0)
    goto out;
```

First of all we store the pid and `pid` that seen from the namespace of the current task in the `exec_binprm` :

```
old_pid = current->pid;
rcu_read_lock();
old_vpid = task_pid_nr_ns(current, task_active_pid_ns(current->parent));
rcu_read_unlock();
```

and call the:

```
search_binary_handler(bprm);
```

function. This function goes through the list of handlers that contains different binary formats. Currently the Linux kernel supports following binary formats:

- `binfmt_script` - support for interpreted scripts that are starts from the #! line;
- `binfmt_misc` - support different binary formats, according to runtime configuration of the Linux kernel;
- `binfmt_elf` - support elf format;
- `binfmt_aout` - support a.out format;
- `binfmt_flat` - support for flat format;
- `binfmt_elf_fdpic` - Support for elf FDPIC binaries;
- `binfmt_em86` - support for Intel elf binaries running on Alpha machines.

So, the `search_binary_handler` tries to call the `load_binary` function and pass `linux_binprm` to it. If the binary handler supports the given executable file format, it starts to prepare the executable binary for execution:

```
int search_binary_handler(struct linux_binprm *bprm)
{
    ...
    ...
    ...
    list_for_each_entry(fmt, &formats, lh) {
        retval = fmt->load_binary(bprm);
        if (retval < 0 && !bprm->mm) {
            force_sigsegv(SIGSEGV, current);
            return retval;
        }
    }

    return retval;
```

Where the `load_binary` for example for the elf checks the magic number (each `elf` binary file contains magic number in the header) in the `linux_bprm` buffer (remember that we read first `128` bytes from the executable binary file): and exit if it is not `elf` binary:

```
static int load_elf_binary(struct linux_binprm *bprm)
{
    ...
```

```
    ...
    ...
    loc->elf_ex = *((struct elfhdr *)bprm->buf);

    if (memcmp(elf_ex.e_ident, ELFMAG, SELFMAG) != 0)
        goto out;
```

If the given executable file is in `elf` format, the `load_elf_binary` continues to execute. The `load_elf_binary` does many different things to prepare on execution executable file. For example it checks the architecture and type of the executable file:

```
if (loc->elf_ex.e_type != ET_EXEC && loc->elf_ex.e_type != ET_DYN)
    goto out;
if (!elf_check_arch(&loc->elf_ex))
    goto out;
```

and exit if there is wrong architecture and executable file non executable non shared. Tries to load the `program header table` :

```
elf_phdata = load_elf_phdrs(&loc->elf_ex, bprm->file);
if (!elf_phdata)
    goto out;
```

that describes segments. Read the `program interpreter` and libraries that linked with the our executable binary file from disk and load it to memory. The `program interpreter` specified in the `.interp` section of the executable file and as you can read in the part that describes Linkers it is - `/lib64/ld-linux-x86-64.so.2` for the `x86_64` . It setups the stack and map `elf` binary into the correct location in memory. It maps the bss and the brk sections and does many many other different things to prepare executable file to execute.

In the end of the execution of the `load_elf_binary` we call the `start_thread` function and pass three arguments to it:

```
    start_thread(regs, elf_entry, bprm->p);
    retval = 0;
out:
    kfree(loc);
out_ret:
    return retval;
```

These arguments are:

- Set of registers for the new task;
- Address of the entry point of the new task;
- Address of the top of the stack for the new task.

As we can understand from the function's name, it starts new thread, but it is not so. The `start_thread` function just prepares new task's registers to be ready to run. Let's look on the implementation of this function:

```
void
start_thread(struct pt_regs *regs, unsigned long new_ip, unsigned long new_sp)
{
        start_thread_common(regs, new_ip, new_sp,
                            __USER_CS, __USER_DS, 0);
}
```

As we can see the `start_thread` function just makes a call of the `start_thread_common` function that will do all for us:

```
static void
start_thread_common(struct pt_regs *regs, unsigned long new_ip,
                    unsigned long new_sp,
                    unsigned int _cs, unsigned int _ss, unsigned int _ds)
```

```
{
        loadsegment(fs, 0);
        loadsegment(es, _ds);
        loadsegment(ds, _ds);
        load_gs_index(0);
        regs->ip                = new_ip;
        regs->sp                = new_sp;
        regs->cs                = _cs;
        regs->ss                = _ss;
        regs->flags             = X86_EFLAGS_IF;
        force_iret();
}
```

The `start_thread_common` function fills `fs` segment register with zero and `es` and `ds` with the value of the data segment register. After this we set new values to the instruction pointer, `cs` segments etc. In the end of the `start_thread_common` function we can see the `force_iret` macro that force a system call return via `iret` instruction. Ok, we prepared new thread to run in userspace and now we can return from the `exec_binprm` and now we are in the `do_execveat_common` again. After the `exec_binprm` will finish its execution we release memory for structures that was allocated before and return.

After we returned from the `execve` system call handler, execution of our program will be started. We can do it, because all context related information already configured for this purpose. As we saw the `execve` system call does not return control to a process, but code, data and other segments of the caller process are just overwritten of the program segments. The exit from our application will be implemented through the `exit` system call.

That's all. From this point our program will be executed.

Conclusion

This is the end of the fourth and last part of the about the system calls concept in the Linux kernel. We saw almost all related stuff to the `system call` concept in these four parts. We started from the understanding of the `system call` concept, we have learned what is it and why do users applications need in this concept. Next we saw how does the Linux handle a system call from a user application. We met two similar concepts to the `system call` concept, they are `vsyscall` and `vDSO` and finally we saw how does Linux kernel run a user program.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- System call
- shell
- bash
- entry point
- C
- environment variables
- file descriptor
- real uid
- virtual file system
- procfs
- sysfs
- inode
- pid

- namespace
- #!
- elf
- a.out
- flat
- Alpha
- FDPIC
- segments
- Linkers
- Processor register
- instruction pointer
- Previous part

How does the `open` system call work

Introduction

This is the fifth part of the chapter that describes system calls mechanism in the Linux kernel. Previous parts of this chapter described this mechanism in general. Now I will try to describe implementation of different system calls in the Linux kernel. Previous parts from this chapter and parts from other chapters of the books describe mostly deep parts of the Linux kernel that are faintly visible or fully invisible from the userspace. But the Linux kernel code is not only about itself. The vast of the Linux kernel code provides ability to our code. Due to the linux kernel our programs can read/write from/to files and don't know anything about sectors, tracks and other parts of a disk structures, we can send data over network and don't build encapsulated network packets by hand and etc.

I don't know how about you, but it is interesting to me not only how an operating system works, but how do my software interacts with it. As you may know, our programs interacts with the kernel through the special mechanism which is called system call. So, I've decided to write series of parts which will describe implementation and behavior of system calls which we are using every day like `read`, `write`, `open`, `close`, `dup` and etc.

I have decided to start from the description of the open system call. if you have written at least one `C` program, you should know that before we are able to read/write or execute other manipulations with a file we need to open it with the `open` function:

```
#include <fcntl.h>
#include <stdio.h>
#include <stdlib.h>
#include <unistd.h>
#include <sys/stat.h>
#include <sys/types.h>

int main(int argc, char *argv) {
        int fd = open("test", O_RDONLY);

        if fd < 0 {
                perror("Opening of the file is failed\n");
        }
        else {
                printf("file sucessfully opened\n");
        }

        close(fd);
        return 0;
}
```

In this case, the open is the function from standard library, but not system call. The standard library will call related system call for us. The `open` call will return a file descriptor which is just an unique number within our process which is associated with the opened file. Now as we opened a file and got file descriptor as result of `open` call, we may start to interact with this file. We can write into, read from it and etc. List of opened file by a process is available via proc filesystem:

```
$ sudo ls /proc/1/fd/

0  10  12  14  16  2   21  23  25  27  29  30  32  34  36  38  4   41  43  45  47  49  50  53  55  58  6   61
63  67  8
1  11  13  15  19  20  22  24  26  28  3   31  33  35  37  39  40  42  44  46  48  5   51  54  57  59  60  62
65  7   9
```

I am not going to describe more details about the `open` routine from the userspace view in this post, but mostly from the kernel side. if you are not very familiar with, you can get more info in the man page.

So let's start.

Definition of the open system call

If you have read the fourth part of the linux-insides book, you should know that system calls are defined with the help of `SYSCALL_DEFINE` macro. So, the `open` system call is not exception.

Definition of the `open` system call is located in the fs/open.c source code file and looks pretty small for the first view:

```
SYSCALL_DEFINE3(open, const char __user *, filename, int, flags, umode_t, mode)
{
    if (force_o_largefile())
        flags |= O_LARGEFILE;

    return do_sys_open(AT_FDCWD, filename, flags, mode);
}
```

As you may guess, the `do_sys_open` function from the same source code file does the main job. But before this function will be called, let's consider the `if` clause from which the implementation of the `open` system call starts:

```
if (force_o_largefile())
    flags |= O_LARGEFILE;
```

Here we apply the `O_LARGEFILE` flag to the flags which were passed to `open` system call in a case when the `force_o_largefile()` will return true. What is `O_LARGEFILE` ? We may read this in the man page for the `open(2)` system call:

> O_LARGEFILE
>
> (LFS) Allow files whose sizes cannot be represented in an off_t (but can be represented in an off64_t) to be opened.

As we may read in the GNU C Library Reference Manual:

> off_t
>
> This is a signed integer type used to represent file sizes. In the GNU C Library, this type is no narrower than int. If the source is compiled with _FILE_OFFSET_BITS == 64 this type is transparently replaced by off64_t.

and

> off64_t
>
> This type is used similar to off_t. The difference is that even on 32 bit machines, where the off_t type would have 32 bits, off64_t has 64 bits and so is able to address files up to 2^63 bytes in length. When compiling with _FILE_OFFSET_BITS == 64 this type is available under the name off_t.

So it is not hard to guess that the `off_t` , `off64_t` and `O_LARGEFILE` are about a file size. In the case of the Linux kernel, the `O_LARGEFILE` is used to disallow opening large files on 32bit systems if the caller didn't specify `O_LARGEFILE` flag during opening of a file. On 64bit systems we force on this flag in open system call. And the `force_o_largefile` macro from the include/linux/fcntl.h linux kernel header file confirms this:

```
#ifndef force_o_largefile
#define force_o_largefile() (BITS_PER_LONG != 32)
#endif
```

This macro may be architecture-specific as for example for IA-64 architecture, but in our case the x86_64 does not provide definition of the `force_o_largefile` and it will be used from include/linux/fcntl.h.

So, as we may see the `force_o_largefile` is just a macro which expands to the `true` value in our case of x86_64 architecture. As we are considering 64-bit architecture, the `force_o_largefile` will be expanded to `true` and the `O_LARGEFILE` flag will be added to the set of flags which were passed to the `open` system call.

Now as we considered meaning of the `O_LARGEFILE` flag and `force_o_largefile` macro, we can proceed to the consideration of the implementation of the `do_sys_open` function. As I wrote above, this function is defined in the same source code file and looks:

```
long do_sys_open(int dfd, const char __user *filename, int flags, umode_t mode)
{
    struct open_flags op;
    int fd = build_open_flags(flags, mode, &op);
    struct filename *tmp;

    if (fd)
        return fd;

    tmp = getname(filename);
    if (IS_ERR(tmp))
        return PTR_ERR(tmp);

    fd = get_unused_fd_flags(flags);
    if (fd >= 0) {
        struct file *f = do_filp_open(dfd, tmp, &op);
        if (IS_ERR(f)) {
            put_unused_fd(fd);
            fd = PTR_ERR(f);
        } else {
            fsnotify_open(f);
            fd_install(fd, f);
        }
    }
    putname(tmp);
    return fd;
}
```

Let's try to understand how the `do_sys_open` works step by step.

open(2) flags

As you know the `open` system call takes set of `flags` as second argument that control opening a file and `mode` as third argument that specifies permission the permissions of a file if it is created. The `do_sys_open` function starts from the call of the `build_open_flags` function which does some checks that set of the given flags is valid and handles different conditions of flags and mode.

Let's look at the implementation of the `build_open_flags`. This function is defined in the same kernel file and takes three arguments:

- flags - flags that control opening of a file;
- mode - permissions for newly created file;

The last argument - `op` is represented with the `open_flags` structure:

```
struct open_flags {
        int open_flag;
        umode_t mode;
        int acc_mode;
        int intent;
        int lookup_flags;
};
```

which is defined in the fs/internal.h header file and as we may see it holds information about flags and access mode for internal kernel purposes. As you already may guess the main goal of the `build_open_flags` function is to fill an instance of this structure.

Implementation of the `build_open_flags` function starts from the definition of local variables and one of them is:

```
int acc_mode = ACC_MODE(flags);
```

This local variable represents access mode and its initial value will be equal to the value of expanded `ACC_MODE` macro. This macro is defined in the include/linux/fs.h and looks pretty interesting:

```
#define ACC_MODE(x) ("\004\002\006\006"[(x)&O_ACCMODE])
#define O_ACCMODE   00000003
```

The `"\004\002\006\006"` is an array of four chars:

```
"\004\002\006\006" == {'\004', '\002', '\006', '\006'}
```

So, the `ACC_MODE` macro just expands to the accession to this array by `[(x) & O_ACCMODE]` index. As we just saw, the `O_ACCMODE` is `00000003`. By applying `x & O_ACCMODE` we will take the two least significant bits which are represents `read`, `write` or `read/write` access modes:

```
#define O_RDONLY        00000000
#define O_WRONLY        00000001
#define O_RDWR          00000002
```

After getting value from the array by the calculated index, the `ACC_MODE` will be expanded to access mode mask of a file which will hold `MAY_WRITE`, `MAY_READ` and other information.

We may see following condition after we have calculated initial access mode:

```
if (flags & (O_CREAT | __O_TMPFILE))
    op->mode = (mode & S_IALLUGO) | S_IFREG;
else
    op->mode = 0;
```

Here we reset permissions in `open_flags` instance if a opened file wasn't temporary and wasn't open for creation. This is because:

> if neither O_CREAT nor O_TMPFILE is specified, then mode is ignored.

In other case if `O_CREAT` or `O_TMPFILE` were passed we canonicalize it to a regular file because a directory should be created with the opendir system call.

At the next step we check that a file is not tried to be opened via fanotify and without the `O_CLOEXEC` flag:

```
flags &= ~FMODE_NONOTIFY & ~O_CLOEXEC;
```

We do this to not leak a file descriptor. By default, the new file descriptor is set to remain open across an `execve` system call, but the `open` system call supports `O_CLOEXEC` flag that can be used to change this default behaviour. So we do this to prevent leaking of a file descriptor when one thread opens a file to set `O_CLOEXEC` flag and in the same time the second process does a fork) + execve) and as you may remember that child will have copies of the parent's set of open file descriptors.

At the next step we check that if our flags contains `O_SYNC` flag, we apply `O_DSYNC` flag too:

```
if (flags & __O_SYNC)
    flags |= O_DSYNC;
```

The `O_SYNC` flag guarantees that the any write call will not return before all data has been transferred to the disk. The `O_DSYNC` is like `O_SYNC` except that there is no requirement to wait for any metadata (like `atime` , `mtime` and etc.) changes will be written. We apply `O_DSYNC` in a case of `__O_SYNC` because it is implemented as `__O_SYNC|O_DSYNC` in the Linux kernel.

After this we must be sure that if a user wants to create temporary file, the flags should contain `O_TMPFILE_MASK` or in other words it should contain or `O_CREAT` or `O_TMPFILE` or both and also it should be writeable:

```
if (flags & __O_TMPFILE) {
    if ((flags & O_TMPFILE_MASK) != O_TMPFILE)
        return -EINVAL;
    if (!(acc_mode & MAY_WRITE))
        return -EINVAL;
} else if (flags & O_PATH) {
           flags &= O_DIRECTORY | O_NOFOLLOW | O_PATH;
        acc_mode = 0;
}
```

as it is written in in the manual page:

> O_TMPFILE must be specified with one of O_RDWR or O_WRONLY

If we didn't pass `O_TMPFILE` for creation of a temporary file, we check the `O_PATH` flag at the next condition. The `O_PATH` flag allows us to obtain a file descriptor that may be used for two following purposes:

- to indicate a location in the filesystem tree;
- to perform operations that act purely at the file descriptor level.

So, in this case the file itself is not opened, but operations like `dup` , `fcntl` and other can be used. So, if all file content related operations like `read` , `write` and other are permitted, only `O_DIRECTORY | O_NOFOLLOW | O_PATH` flags can be used. We have finished with flags for this moment in the `build_open_flags` for this moment and we may fill our `open_flags->open_flag` with them:

```
op->open_flag = flags;
```

Now we have filled `open_flag` field which represents flags that will control opening of a file and `mode` that will represent `umask` of a new file if we open file for creation. There are still to fill last flags in the our `open_flags` structure. The next is `op->acc_mode` which represents access mode to a opened file. We already filled the `acc_mode` local variable with the initial value at the beginning of the `build_open_flags` and now we check last two flags related to access mode:

```
if (flags & O_TRUNC)
        acc_mode |= MAY_WRITE;
if (flags & O_APPEND)
    acc_mode |= MAY_APPEND;
op->acc_mode = acc_mode;
```

These flags are - `O_TRUNC` that will truncate an opened file to length `0` if it existed before we open it and the `O_APPEND` flag allows to open a file in `append mode` . So the opened file will be appended during write but not overwritten.

The next field of the `open_flags` structure is - `intent` . It allows us to know about our intention or in other words what do we really want to do with file, open it, create, rename it or something else. So we set it to zero if our flags contains the `O_PATH` flag as we can't do anything related to a file content with this flag:

```
op->intent = flags & O_PATH ? 0 : LOOKUP_OPEN;
```

or just to `LOOKUP_OPEN` intention. Additionally we set `LOOKUP_CREATE` intention if we want to create new file and to be sure that a file didn't exist before with `O_EXCL` flag:

```
if (flags & O_CREAT) {
    op->intent |= LOOKUP_CREATE;
    if (flags & O_EXCL)
        op->intent |= LOOKUP_EXCL;
}
```

The last flag of the `open_flags` structure is the `lookup_flags` :

```
if (flags & O_DIRECTORY)
    lookup_flags |= LOOKUP_DIRECTORY;
if (!(flags & O_NOFOLLOW))
    lookup_flags |= LOOKUP_FOLLOW;
op->lookup_flags = lookup_flags;

return 0;
```

We fill it with `LOOKUP_DIRECTORY` if we want to open a directory and `LOOKUP_FOLLOW` if we don't want to follow (open) symlink. That's all. It is the end of the `build_open_flags` function. The `open_flags` structure is filled with modes and flags for a file opening and we can return back to the `do_sys_open` .

Actual opening of a file

At the next step after `build_open_flags` function is finished and we have formed flags and modes for our file we should get the `filename` structure with the help of the `getname` function by name of a file which was passed to the `open` system call:

```
tmp = getname(filename);
if (IS_ERR(tmp))
    return PTR_ERR(tmp);
```

The `getname` function is defined in the fs/namei.c source code file and looks:

```
struct filename *
getname(const char __user * filename)
{
        return getname_flags(filename, 0, NULL);
}
```

So, it just calls the `getname_flags` function and returns its result. The main goal of the `getname_flags` function is to copy a file path given from userland to kernel space. The `filename` structure is defined in the include/linux/fs.h linux kernel header file and contains following fields:

- name - pointer to a file path in kernel space;
- uptr - original pointer from userland;
- aname - filename from audit context;
- refcnt - reference counter;
- iname - a filename in a case when it will be less than `PATH_MAX` .

As I already wrote above, the main goal of the `getname_flags` function is to copy name of a file which was passed to the `open` system call from user space to kernel space with the strncpy_from_user function. The next step after a filename will be copied to kernel space is getting of new non-busy file descriptor:

```
fd = get_unused_fd_flags(flags);
```

The `get_unused_fd_flags` function takes table of open files of the current process, minimum (`0`) and maximum (`RLIMIT_NOFILE`) possible number of a file descriptor in the system and flags that we have passed to the `open` system call and allocates file descriptor and mark it busy in the file descriptor table of the current process. The `get_unused_fd_flags` function sets or clears the `O_CLOEXEC` flag depends on its state in the passed flags.

The last and main step in the `do_sys_open` is the `do_filp_open` function:

```
struct file *f = do_filp_open(dfd, tmp, &op);

if (IS_ERR(f)) {
    put_unused_fd(fd);
    fd = PTR_ERR(f);
} else {
    fsnotify_open(f);
    fd_install(fd, f);
}
```

The main goal of this function is to resolve given path name into `file` structure which represents an opened file of a process. If something going wrong and execution of the `do_filp_open` function will be failed, we should free new file descriptor with the `put_unused_fd` or in other way the `file` structure returned by the `do_filp_open` will be stored in the file descriptor table of the current process.

Now let's take a short look at the implementation of the `do_filp_open` function. This function is defined in the fs/namei.c linux kernel source code file and starts from initialization of the `nameidata` structure. This structure will provide a link to a file inode. Actually this is one of the main point of the `do_filp_open` function to acquire an `inode` by the filename given to `open` system call. After the `nameidata` structure will be initialized, the `path_openat` function will be called:

```
filp = path_openat(&nd, op, flags | LOOKUP_RCU);

if (unlikely(filp == ERR_PTR(-ECHILD)))
    filp = path_openat(&nd, op, flags);
if (unlikely(filp == ERR_PTR(-ESTALE)))
    filp = path_openat(&nd, op, flags | LOOKUP_REVAL);
```

Note that it is called three times. Actually, the Linux kernel will open the file in RCU mode. This is the most efficient way to open a file. If this try will be failed, the kernel enters the normal mode. The third call is relatively rare, only in the nfs file system is likely to be used. The `path_openat` function executes `path lookup` or in other words it tries to find a `dentry` (what the Linux kernel uses to keep track of the hierarchy of files in directories) corresponding to a path.

The `path_openat` function starts from the call of the `get_empty_flip()` function that allocates a new `file` structure with some additional checks like do we exceed amount of opened files in the system or not and etc. After we have got allocated new `file` structure we call the `do_tmpfile` or `do_o_path` functions in a case if we have passed `O_TMPFILE | O_CREATE` or `O_PATH` flags during call of the `open` system call. These both cases are quite specific, so let's consider quite usual case when we want to open already existed file and want to read/write from/to it.

In this case the `path_init` function will be called. This function performs some preporatory work before actual path lookup. This includes search of start position of path traversal and its metadata like `inode` of the path, `dentry inode` and etc. This can be `root` directory - `/` or current directory as in our case, because we use `AT_CWD` as starting point (see call of the `do_sys_open` at the beginning of the post).

The next step after the `path_init` is the loop which executes the `link_path_walk` and `do_last` . The first function executes name resolution or in other words this function starts process of walking along a given path. It handles everything step by step except the last component of a file path. This handling includes checking of a permissions and getting a file component. As a file component is gotten, it is passed to `walk_component` that updates current directory entry from the `dcache` or asks underlying filesystem. This repeats before all path's components will not be handled in such way. After the `link_path_walk` will be executed, the `do_last` function will populate a `file` structure based on the result of the `link_path_walk`. As we reached last component of the given file path the `vfs_open` function from the `do_last` will be called.

This function is defined in the fs/open.c linux kernel source code file and the main goal of this function is to call an `open` operation of underlying filesystem.

That's all for now. We didn't consider **full** implementation of the `open` system call. We skip some parts like handling case when we want to open a file from other filesystem with different mount point, resolving symlinks and etc., but it should be not so hard to follow this stuff. This stuff does not included in **generic** implementation of open system call and depends on underlying filesystem. If you are interested in, you may lookup the `file_operations.open` callback function for a certain filesystem.

Conclusion

This is the end of the fifth part of the implementation of different system calls in the Linux kernel. If you have questions or suggestions, ping me on twitter 0xAX, drop me an email, or just create an issue. In the next part, we will continue to dive into system calls in the Linux kernel and see the implementation of the read system call.

Please note that English is not my first language and I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- system call
- open
- file descriptor
- proc
- GNU C Library Reference Manual
- IA-64
- x86_64
- opendir
- fanotify
- fork)
- execve)
- symlink
- audit
- inode
- RCU
- read
- previous part

Limits on resources in Linux

Each process in the system uses certain amount of different resources like files, CPU time, memory and so on.

Such resources are not infinite and each process and we should have an instrument to manage it. Sometimes it is useful to know current limits for a certain resource or to change it's value. In this post we will consider such instruments that allow us to get information about limits for a process and increase or decrease such limits.

We will start from userspace view and then we will look how it is implemented in the Linux kernel.

There are three main fundamental system calls to manage resource limit for a process:

- `getrlimit`
- `setrlimit`
- `prlimit`

The first two allows a process to read and set limits on a system resource. The last one is extension for previous functions. The `prlimit` allows to set and read the resource limits of a process specified by PID. Definitions of these functions looks:

The `getrlimit` is:

```
int getrlimit(int resource, struct rlimit *rlim);
```

The `setrlimit` is:

```
int setrlimit(int resource, const struct rlimit *rlim);
```

And the definition of the `prlimit` is:

```
int prlimit(pid_t pid, int resource, const struct rlimit *new_limit,
            struct rlimit *old_limit);
```

In the first two cases, functions takes two parameters:

- `resource` - represents resource type (we will see available types later);
- `rlim` - combination of `soft` and `hard` limits.

There are two types of limits:

- `soft`
- `hard`

The first provides actual limit for a resource of a process. The second is a ceiling value of a `soft` limit and can be set only by superuser. So, `soft` limit can never exceed related `hard` limit.

Both these values are combined in the `rlimit` structure:

```
struct rlimit {
    rlim_t rlim_cur;
    rlim_t rlim_max;
};
```

The last one function looks a little bit complex and takes `4` arguments. Besides `resource` argument, it takes:

- `pid` - specifies an ID of a process on which the `prlimit` should be executed;

* `new_limit` - provides new limits values if it is not `NULL` ;
* `old_limit` - current `soft` and `hard` limits will be placed here if it is not `NULL` .

Exactly `prlimit` function is used by ulimit util. We can verify this with the help of strace util.

For example:

```
~$ strace ulimit -s 2>&1 | grep rl

prlimit64(0, RLIMIT_NPROC, NULL, {rlim_cur=63727, rlim_max=63727}) = 0
prlimit64(0, RLIMIT_NOFILE, NULL, {rlim_cur=1024, rlim_max=4*1024}) = 0
prlimit64(0, RLIMIT_STACK, NULL, {rlim_cur=8192*1024, rlim_max=RLIM64_INFINITY}) = 0
```

Here we can see `prlimit64` , but not the `prlimit` . The fact is that we see underlying system call here instead of library call.

Now let's look at list of available resources:

Resource	Description
RLIMIT_CPU	CPU time limit given in seconds
RLIMIT_FSIZE	the maximum size of files that a process may create
RLIMIT_DATA	the maximum size of the process's data segment
RLIMIT_STACK	the maximum size of the process stack in bytes
RLIMIT_CORE	the maximum size of a core file.
RLIMIT_RSS	the number of bytes that can be allocated for a process in RAM
RLIMIT_NPROC	the maximum number of processes that can be created by a user
RLIMIT_NOFILE	the maximum number of a file descriptor that can be opened by a process
RLIMIT_MEMLOCK	the maximum number of bytes of memory that may be locked into RAM by mlock.
RLIMIT_AS	the maximum size of virtual memory in bytes.
RLIMIT_LOCKS	the maximum number flock and locking related fcntl calls
RLIMIT_SIGPENDING	maximum number of signals that may be queued for a user of the calling process
RLIMIT_MSGQUEUE	the number of bytes that can be allocated for POSIX message queues
RLIMIT_NICE	the maximum nice value that can be set by a process
RLIMIT_RTPRIO	maximum real-time priority value
RLIMIT_RTTIME	maximum number of microseconds that a process may be scheduled under real-time scheduling policy without making blocking system call

If you're looking into source code of open source projects, you will note that reading or updating of a resource limit is quite widely used operation.

For example: systemd

```
/* Don't limit the coredump size */
(void) setrlimit(RLIMIT_CORE, &RLIMIT_MAKE_CONST(RLIM_INFINITY));
```

Or haproxy:

```
getrlimit(RLIMIT_NOFILE, &limit);
if (limit.rlim_cur < global.maxsock) {
    Warning("[%s.main()] FD limit (%d) too low for maxconn=%d/maxsock=%d. Please raise 'ulimit-n' to %d or more
```

```
 to avoid any trouble.\n",
        argv[0], (int)limit.rlim_cur, global.maxconn, global.maxsock, global.maxsock);
}
```

We've just saw a little bit about resources limits related stuff in the userspace, now let's look at the same system calls in the Linux kernel.

Limits on resource in the Linux kernel

Both implementation of `getrlimit` system call and `setrlimit` looks similar. Both they execute `do_prlimit` function that is core implementation of the `prlimit` system call and copy from/to given `rlimit` from/to userspace:

The `getrlimit` :

```
SYSCALL_DEFINE2(getrlimit, unsigned int, resource, struct rlimit __user *, rlim)
{
    struct rlimit value;
    int ret;

    ret = do_prlimit(current, resource, NULL, &value);
    if (!ret)
        ret = copy_to_user(rlim, &value, sizeof(*rlim)) ? -EFAULT : 0;

    return ret;
}
```

and `setrlimit` :

```
SYSCALL_DEFINE2(setrlimit, unsigned int, resource, struct rlimit __user *, rlim)
{
    struct rlimit new_rlim;

    if (copy_from_user(&new_rlim, rlim, sizeof(*rlim)))
        return -EFAULT;
    return do_prlimit(current, resource, &new_rlim, NULL);
}
```

Implementations of these system calls are defined in the kernel/sys.c kernel source code file.

First of all the `do_prlimit` function executes a check that the given resource is valid:

```
if (resource >= RLIM_NLIMITS)
    return -EINVAL;
```

and in a failure case returns `-EINVAL` error. After this check will pass successfully and new limits was passed as non `NULL` value, two following checks:

```
if (new_rlim) {
    if (new_rlim->rlim_cur > new_rlim->rlim_max)
        return -EINVAL;
    if (resource == RLIMIT_NOFILE &&
            new_rlim->rlim_max > sysctl_nr_open)
        return -EPERM;
}
```

check that the given `soft` limit does not exceed `hard` limit and in a case when the given resource is the maximum number of a file descriptors that hard limit is not greater than `sysctl_nr_open` value. The value of the `sysctl_nr_open` can be found via procfs:

```
~$ cat /proc/sys/fs/nr_open
1048576
```

After all of these checks we lock `tasklist` to be sure that signal handlers related things will not be destroyed while we updating limits for a given resource:

```
read_lock(&tasklist_lock);
...
...
...
read_unlock(&tasklist_lock);
```

We need to do this because `prlimit` system call allows us to update limits of another task by the given pid. As task list is locked, we take the `rlimit` instance that is responsible for the given resource limit of the given process:

```
rlim = tsk->signal->rlim + resource;
```

where the `tsk->signal->rlim` is just array of `struct rlimit` that represents certain resources. And if the `new_rlim` is not `NULL` we just update its value. If `old_rlim` is not `NULL` we fill it:

```
if (old_rlim)
    *old_rlim = *rlim;
```

That's all.

Conclusion

This is the end of the second part that describes implementation of the system calls in the Linux kernel. If you have questions or suggestions, ping me on Twitter 0xAX, drop me an email, or just create an issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you find any mistakes please send me PR to linux-insides.

Links

- system calls
- PID
- ulimit
- strace
- POSIX message queues

Timers and time management

This chapter describes timers and time management related concepts in the linux kernel.

- Introduction - An introduction to the timers in the Linux kernel.
- Introduction to the clocksource framework - Describes `clocksource` framework in the Linux kernel.
- The tick broadcast framework and dyntick - Describes tick broadcast framework and dyntick concept.
- Introduction to timers - Describes timers in the Linux kernel.
- Introduction to the clockevents framework - Describes yet another clock/time management related framework : `clockevents`.
- x86 related clock sources - Describes `x86_64` related clock sources.
- Time related system calls in the Linux kernel - Describes time related system calls.

Timers and time management in the Linux kernel. Part 1.

Introduction

This is yet another post that opens a new chapter in the linux-insides book. The previous part described system call concepts, and now it's time to start new chapter. As one might understand from the title, this chapter will be devoted to the `timers` and `time management` in the Linux kernel. The choice of topic for the current chapter is not accidental. Timers (and generally, time management) are very important and widely used in the Linux kernel. The Linux kernel uses timers for various tasks, for example different timeouts in the TCP implementation, the kernel knowing current time, scheduling asynchronous functions, next event interrupt scheduling and many many more.

So, we will start to learn implementation of the different time management related stuff in this part. We will see different types of timers and how different Linux kernel subsystems use them. As always, we will start from the earliest part of the Linux kernel and go through the initialization process of the Linux kernel. We already did it in the special chapter which describes the initialization process of the Linux kernel, but as you may remember we missed some things there. And one of them is the initialization of timers.

Let's start.

Initialization of non-standard PC hardware clock

After the Linux kernel was decompressed (more about this you can read in the Kernel decompression part) the architecture non-specific code starts to work in the init/main.c source code file. After initialization of the lock validator, initialization of cgroups and setting canary value we can see the call of the `setup_arch` function.

As you may remember, this function (defined in the arch/x86/kernel/setup.c) prepares/initializes architecture-specific stuff (for example it reserves a place for bss section, reserves a place for initrd, parses kernel command line, and many, many other things). Besides this, we can find some time management related functions there.

The first is:

```
x86_init.timers.wallclock_init();
```

We already saw `x86_init` structure in the chapter that describes initialization of the Linux kernel. This structure contains pointers to the default setup functions for the different platforms like Intel MID, Intel CE4100, etc. The `x86_init` structure is defined in the arch/x86/kernel/x86_init.c, and as you can see it determines standard PC hardware by default.

As we can see, the `x86_init` structure has the `x86_init_ops` type that provides a set of functions for platform specific setup like reserving standard resources, platform specific memory setup, initialization of interrupt handlers, etc. This structure looks like:

```
struct x86_init_ops {
    struct x86_init_resources       resources;
    struct x86_init_mpparse         mpparse;
    struct x86_init_irqs            irqs;
    struct x86_init_oem             oem;
    struct x86_init_paging          paging;
    struct x86_init_timers          timers;
    struct x86_init_iommu           iommu;
    struct x86_init_pci             pci;
};
```

Note the `timers` field that has the `x86_init_timers` type. We can understand by its name that this field is related to time management and timers. `x86_init_timers` contains four fields which are all functions that returns pointer on void:

- `setup_percpu_clockev` - set up the per cpu clock event device for the boot cpu;
- `tsc_pre_init` - platform function called before TSC init;
- `timer_init` - initialize the platform timer;
- `wallclock_init` - initialize the wallclock device.

So, as we already know, in our case the `wallclock_init` executes initialization of the wallclock device. If we look on the `x86_init` structure, we see that `wallclock_init` points to the `x86_init_noop` :

```
struct x86_init_ops x86_init __initdata = {
    ...
    ...
    ...
    .timers = {
        .wallclock_init            = x86_init_noop,
    },
    ...
    ...
    ...
}
```

Where the `x86_init_noop` is just a function that does nothing:

```
void __cpuinit x86_init_noop(void) { }
```

for the standard PC hardware. Actually, the `wallclock_init` function is used in the Intel MID platform. Initialization of the `x86_init.timers.wallclock_init` is located in the arch/x86/platform/intel-mid/intel-mid.c source code file in the `x86_intel_mid_early_setup` function:

```
void __init x86_intel_mid_early_setup(void)
{
    ...
    ...
    ...
    x86_init.timers.wallclock_init = intel_mid_rtc_init;
    ...
    ...
    ...
}
```

Implementation of the `intel_mid_rtc_init` function is in the arch/x86/platform/intel-mid/intel_mid_vrtc.c source code file and looks pretty simple. First of all, this function parses Simple Firmware Interface M-Real-Time-Clock table for getting such devices to the `sfi_mrtc_array` array and initialization of the `set_time` and `get_time` functions:

```
void __init intel_mid_rtc_init(void)
{
    unsigned long vrtc_paddr;

    sfi_table_parse(SFI_SIG_MRTC, NULL, NULL, sfi_parse_mrtc);

    vrtc_paddr = sfi_mrtc_array[0].phys_addr;
    if (!sfi_mrtc_num || !vrtc_paddr)
        return;

    vrtc_virt_base = (void __iomem *)set_fixmap_offset_nocache(FIX_LNW_VRTC,
                                vrtc_paddr);

    x86_platform.get_wallclock = vrtc_get_time;
```

```
    x86_platform.set_wallclock = vrtc_set_mmss;
}
```

That's all, after this a device based on `Intel MID` will be able to get time from the hardware clock. As I already wrote, the standard PC x86_64 architecture does not support `x86_init_noop` and just do nothing during call of this function. We just saw initialization of the real time clock for the Intel MID architecture, now it's time to return to the general `x86_64` architecture and will look on the time management related stuff there.

Acquainted with jiffies

If we return to the `setup_arch` function (which is located, as you remember, in the arch/x86/kernel/setup.c source code file), we see the next call of the time management related function:

```
register_refined_jiffies(CLOCK_TICK_RATE);
```

Before we look at the implementation of this function, we must know about jiffy. As we can read on wikipedia:

```
Jiffy is an informal term for any unspecified short period of time
```

This definition is very similar to the `jiffy` in the Linux kernel. There is global variable with the `jiffies` which holds the number of ticks that have occurred since the system booted. The Linux kernel sets this variable to zero:

```
extern unsigned long volatile __jiffy_data jiffies;
```

during initialization process. This global variable will be increased each time during timer interrupt. Besides this, near the `jiffies` variable we can see the definition of the similar variable

```
extern u64 jiffies_64;
```

Actually, only one of these variables is in use in the Linux kernel, and it depends on the processor type. For the x86_64 it will be `u64` use and for the x86 it's `unsigned long`. We see this looking at the arch/x86/kernel/vmlinux.lds.S linker script:

```
#ifdef CONFIG_X86_32
...
jiffies = jiffies_64;
...
#else
...
jiffies_64 = jiffies;
...
#endif
```

In the case of `x86_32` the `jiffies` will be the lower `32` bits of the `jiffies_64` variable. Schematically, we can imagine it as follows

```
                    jiffies_64
+-----------------------------------------------------+
|                       |                             |
|                       |                             |
|                       |       jiffies on `x86_32`   |
|                       |                             |
|                       |                             |
+-----------------------------------------------------+
63                     31                             0
```

Now we know a little theory about `jiffies` and can return to our function. There is no architecture-specific implementation for our function - the `register_refined_jiffies`. This function is located in the generic kernel code - kernel/time/jiffies.c source code file. Main point of the `register_refined_jiffies` is registration of the jiffy `clocksource`. Before we look on the implementation of the `register_refined_jiffies` function, we must know what `clocksource` is. As we can read in the comments:

```
The `clocksource` is hardware abstraction for a free-running counter.
```

I'm not sure about you, but that description didn't give a good understanding about the `clocksource` concept. Let's try to understand what is it, but we will not go deeper because this topic will be described in a separate part in much more detail. The main point of the `clocksource` is timekeeping abstraction or in very simple words - it provides a time value to the kernel. We already know about the `jiffies` interface that represents number of ticks that have occurred since the system booted. It is represented by a global variable in the Linux kernel and increases each timer interrupt. The Linux kernel can use `jiffies` for time measurement. So why do we need in separate context like the `clocksource`? Actually, different hardware devices provide different clock sources that are varied in their capabilities. The availability of more precise techniques for time intervals measurement is hardware-dependent.

For example `x86` has on-chip a 64-bit counter that is called Time Stamp Counter and its frequency can be equal to processor frequency. Or for example the High Precision Event Timer, that consists of a `64-bit` counter of at least `10 MHz` frequency. Two different timers and they are both for `x86`. If we will add timers from other architectures, this only makes this problem more complex. The Linux kernel provides the `clocksource` concept to solve the problem.

The clocksource concept is represented by the `clocksource` structure in the Linux kernel. This structure is defined in the include/linux/clocksource.h header file and contains a couple of fields that describe a time counter. For example, it contains - `name` field which is the name of a counter, `flags` field that describes different properties of a counter, pointers to the `suspend` and `resume` functions, and many more.

Let's look at the `clocksource` structure for jiffies that is defined in the kernel/time/jiffies.c source code file:

```
static struct clocksource clocksource_jiffies = {
    .name        = "jiffies",
    .rating        = 1,
    .read        = jiffies_read,
    .mask        = 0xffffffff,
    .mult        = NSEC_PER_JIFFY << JIFFIES_SHIFT,
    .shift        = JIFFIES_SHIFT,
    .max_cycles    = 10,
};
```

We can see the definition of the default name here - `jiffies`. The next is the `rating` field, which allows the best registered clock source to be chosen by the clock source management code available for the specified hardware. The `rating` may have following value:

- `1-99` - Only available for bootup and testing purposes;
- `100-199` - Functional for real use, but not desired.
- `200-299` - A correct and usable clocksource.
- `300-399` - A reasonably fast and accurate clocksource.
- `400-499` - The ideal clocksource. A must-use where available;

For example, rating of the time stamp counter is `300`, but rating of the high precision event timer is `250`. The next field is `read` - it is pointer to the function that allows it to read clocksource's cycle value; or in other words, it just returns `jiffies` variable with `cycle_t` type:

```
static cycle_t jiffies_read(struct clocksource *cs)
{
        return (cycle_t) jiffies;
```

```
}
```

that is just 64-bit unsigned type:

```
typedef u64 cycle_t;
```

The next field is the `mask` value, which ensures that subtraction between counters values from non `64 bit` counters do not need special overflow logic. In our case the mask is `0xffffffff` and it is `32` bits. This means that `jiffy` wraps around to zero after `42` seconds:

```
>>> 0xffffffff
4294967295
# 42 nanoseconds
>>> 42 * pow(10, -9)
4.2000000000000006e-08
# 43 nanoseconds
>>> 43 * pow(10, -9)
4.3e-08
```

The next two fields `mult` and `shift` are used to convert the clocksource's period to nanoseconds per cycle. When the kernel calls the `clocksource.read` function, this function returns a value in `machine` time units represented with `cycle_t` data type that we saw just now. To convert this return value to nanoseconds we need these two fields: `mult` and `shift`. The `clocksource` provides the `clocksource_cyc2ns` function that will do it for us with the following expression:

```
((u64) cycles * mult) >> shift;
```

As we can see the `mult` field is equal:

```
NSEC_PER_JIFFY << JIFFIES_SHIFT

#define NSEC_PER_JIFFY  ((NSEC_PER_SEC+HZ/2)/HZ)
#define NSEC_PER_SEC    1000000000L
```

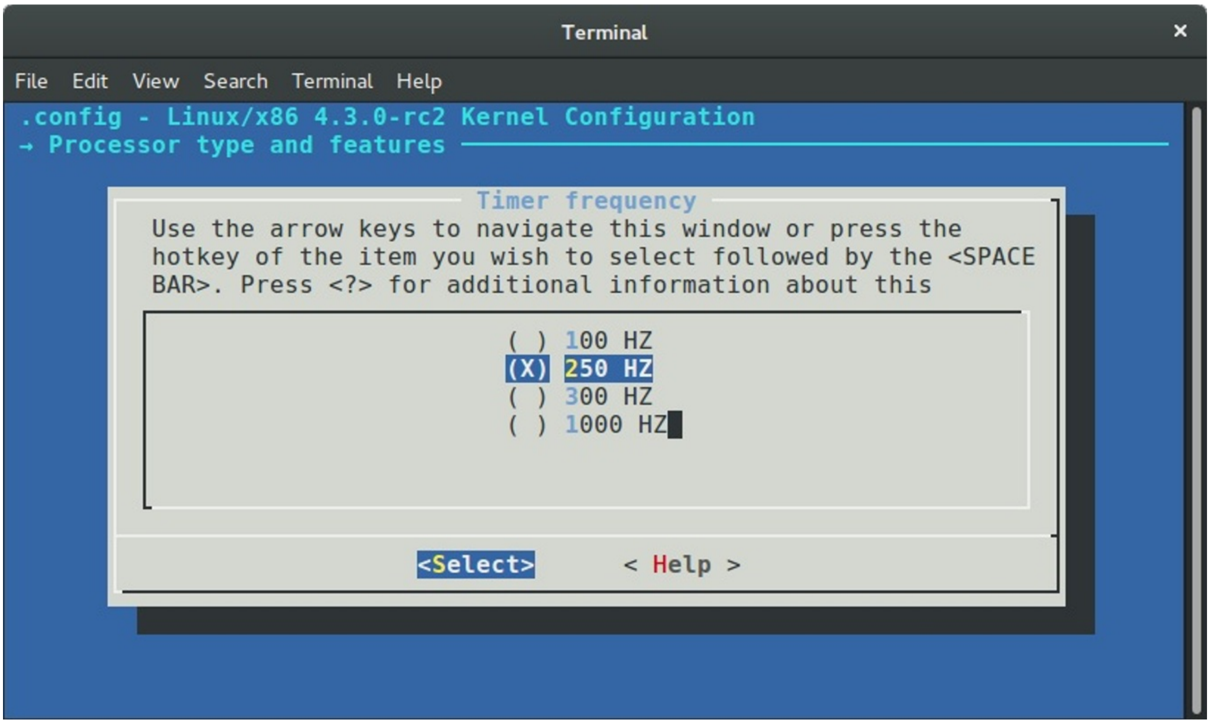
by default, and the `shift` is

```
#if HZ < 34
  #define JIFFIES_SHIFT   6
#elif HZ < 67
  #define JIFFIES_SHIFT   7
#else
  #define JIFFIES_SHIFT   8
#endif
```

The `jiffies` clock source uses the `NSEC_PER_JIFFY` multiplier conversion to specify the nanosecond over cycle ratio. Note that values of the `JIFFIES_SHIFT` and `NSEC_PER_JIFFY` depend on `HZ` value. The `HZ` represents the frequency of the system timer. This macro defined in the include/asm-generic/param.h and depends on the `CONFIG_HZ` kernel configuration option. The value of `HZ` differs for each supported architecture, but for `x86` it's defined like:

```
#define HZ        CONFIG_HZ
```

Where `CONFIG_HZ` can be one of the following values:



This means that in our case the timer interrupt frequency is `250 HZ` or occurs `250` times per second or one timer interrupt each `4ms` .

The last field that we can see in the definition of the `clocksource_jiffies` structure is the - `max_cycles` that holds the maximum cycle value that can safely be multiplied without potentially causing an overflow.

Ok, we just saw definition of the `clocksource_jiffies` structure, also we know a little about `jiffies` and `clocksource` , now it is time to get back to the implementation of the our function. In the beginning of this part we have stopped on the call of the:

```
register_refined_jiffies(CLOCK_TICK_RATE);
```

function from the arch/x86/kernel/setup.c source code file.

As I already wrote, the main purpose of the `register_refined_jiffies` function is to register `refined_jiffies` clocksource. We already saw the `clocksource_jiffies` structure represents standard `jiffies` clock source. Now, if you look in the kernel/time/jiffies.c source code file, you will find yet another clock source definition:

```
struct clocksource refined_jiffies;
```

There is one difference between `refined_jiffies` and `clocksource_jiffies` : The standard `jiffies` based clock source is the lowest common denominator clock source which should function on all systems. As we already know, the `jiffies` global variable will be increased during each timer interrupt. This means the that standard `jiffies` based clock source has the same resolution as the timer interrupt frequency. From this we can understand that standard `jiffies` based clock source may suffer from inaccuracies. The `refined_jiffies` uses `CLOCK_TICK_RATE` as the base of `jiffies` shift.

Let's look at the implementation of this function. First of all, we can see that the `refined_jiffies` clock source based on the `clocksource_jiffies` structure:

```
int register_refined_jiffies(long cycles_per_second)
{
    u64 nsec_per_tick, shift_hz;
    long cycles_per_tick;

    refined_jiffies = clocksource_jiffies;
```

```
    refined_jiffies.name = "refined-jiffies";
    refined_jiffies.rating++;
    ...
    ...
    ...
```

Here we can see that we update the name of the `refined_jiffies` to `refined-jiffies` and increase the rating of this structure. As you remember, the `clocksource_jiffies` has rating - `1`, so our `refined_jiffies` clocksource will have rating - `2`. This means that the `refined_jiffies` will be the best selection for clock source management code.

In the next step we need to calculate number of cycles per one tick:

```
cycles_per_tick = (cycles_per_second + HZ/2)/HZ;
```

Note that we have used `NSEC_PER_SEC` macro as the base of the standard `jiffies` multiplier. Here we are using the `cycles_per_second` which is the first parameter of the `register_refined_jiffies` function. We've passed the `CLOCK_TICK_RATE` macro to the `register_refined_jiffies` function. This macro is defined in the arch/x86/include/asm/timex.h header file and expands to the:

```
#define CLOCK_TICK_RATE         PIT_TICK_RATE
```

where the `PIT_TICK_RATE` macro expands to the frequency of the Intel 8253:

```
#define PIT_TICK_RATE 1193182ul
```

After this we calculate `shift_hz` for the `register_refined_jiffies` that will store `hz << 8` or in other words frequency of the system timer. We shift left the `cycles_per_second` or frequency of the programmable interval timer on `8` in order to get extra accuracy:

```
shift_hz = (u64)cycles_per_second << 8;
shift_hz += cycles_per_tick/2;
do_div(shift_hz, cycles_per_tick);
```

In the next step we calculate the number of seconds per one tick by shifting left the `NSEC_PER_SEC` on `8` too as we did it with the `shift_hz` and do the same calculation as before:

```
nsec_per_tick = (u64)NSEC_PER_SEC << 8;
nsec_per_tick += (u32)shift_hz/2;
do_div(nsec_per_tick, (u32)shift_hz);
```

```
refined_jiffies.mult = ((u32)nsec_per_tick) << JIFFIES_SHIFT;
```

In the end of the `register_refined_jiffies` function we register new clock source with the `__clocksource_register` function that is defined in the include/linux/clocksource.h header file and return:

```
__clocksource_register(&refined_jiffies);
return 0;
```

The clock source management code provides the API for clock source registration and selection. As we can see, clock sources are registered by calling the `__clocksource_register` function during kernel initialization or from a kernel module. During registration, the clock source management code will choose the best clock source available in the system using the `clocksource.rating` field which we already saw when we initialized `clocksource` structure for `jiffies`.

Using the jiffies

We just saw initialization of two `jiffies` based clock sources in the previous paragraph:

- standard `jiffies` based clock source;
- refined `jiffies` based clock source;

Don't worry if you don't understand the calculations here. They look frightening at first. Soon, step by step we will learn these things. So, we just saw initialization of `jiffies` based clock sources and also we know that the Linux kernel has the global variable `jiffies` that holds the number of ticks that have occurred since the kernel started to work. Now, let's look how to use it. To use `jiffies` we just can use the `jiffies` global variable by its name or with the call of the `get_jiffies_64` function. This function defined in the kernel/time/jiffies.c source code file and just returns full `64-bit` value of the `jiffies` :

```
u64 get_jiffies_64(void)
{
    unsigned long seq;
    u64 ret;

    do {
        seq = read_seqbegin(&jiffies_lock);
        ret = jiffies_64;
    } while (read_seqretry(&jiffies_lock, seq));
    return ret;
}
EXPORT_SYMBOL(get_jiffies_64);
```

Note that the `get_jiffies_64` function does not implemented as `jiffies_read` for example:

```
static cycle_t jiffies_read(struct clocksource *cs)
{
    return (cycle_t) jiffies;
}
```

We can see that implementation of the `get_jiffies_64` is more complex. The reading of the `jiffies_64` variable is implemented using seqlocks. Actually this is done for machines that cannot atomically read the full 64-bit values.

If we can access the `jiffies` or the `jiffies_64` variable we can convert it to `human` time units. To get one second we can use following expression:

```
jiffies / HZ
```

So, if we know this, we can get any time units. For example:

```
/* Thirty seconds from now */
jiffies + 30*HZ

/* Two minutes from now */
jiffies + 120*HZ

/* One millisecond from now */
jiffies + HZ / 1000
```

That's all.

Conclusion

This concludes the first part covering time and time management related concepts in the Linux kernel. We first met two concepts and their initialization: `jiffies` and `clocksource`. In the next part we will continue to dive into this interesting theme, and as I already wrote in this part, we will try to understand the insides of these and other time management concepts in the Linux kernel.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- system call
- TCP
- lock validator
- cgroups
- bss
- initrd
- Intel MID
- TSC
- void
- Simple Firmware Interface
- x86_64
- real time clock
- Jiffy
- high precision event timer
- nanoseconds
- Intel 8253
- seqlocks
- cloksource documentation
- Previous chapter

Timers and time management in the Linux kernel. Part 2.

Introduction to the `clocksource` framework

The previous part was the first part in the current chapter that describes timers and time management related stuff in the Linux kernel. We got acquainted with two concepts in the previous part:

- `jiffies`
- `clocksource`

The first is the global variable that is defined in the include/linux/jiffies.h header file and represents the counter that is increased during each timer interrupt. So if we can access this global variable and we know the timer interrupt rate we can convert `jiffies` to the human time units. As we already know the timer interrupt rate represented by the compile-time constant that is called `HZ` in the Linux kernel. The value of `HZ` is equal to the value of the `CONFIG_HZ` kernel configuration option and if we will look into the arch/x86/configs/x86_64_defconfig kernel configuration file, we will see that:

```
CONFIG_HZ_1000=y
```

kernel configuration option is set. This means that value of `CONFIG_HZ` will be `1000` by default for the x86_64 architecture. So, if we divide the value of `jiffies` by the value of `HZ` :

```
jiffies / HZ
```

we will get the amount of seconds that elapsed since the beginning of the moment the Linux kernel started to work or in other words we will get the system uptime. Since `HZ` represents the amount of timer interrupts in a second, we can set a value for some time in the future. For example:

```
/* one minute from now */
unsigned long later = jiffies + 60*HZ;

/* five minutes from now */
unsigned long later = jiffies + 5*60*HZ;
```

This is a very common practice in the Linux kernel. For example, if you will look into the arch/x86/kernel/smpboot.c source code file, you will find the `do_boot_cpu` function. This function boots all processors besides bootstrap processor. You can find a snippet that waits ten seconds for a response from the application processor:

```
if (!boot_error) {
    timeout = jiffies + 10*HZ;
    while (time_before(jiffies, timeout)) {
        ...
        ...
        ...
        udelay(100);
    }
    ...
    ...
    ...
}
```

We assign `jiffies + 10*HZ` value to the `timeout` variable here. As I think you already understood, this means a ten seconds timeout. After this we are entering a loop where we use the `time_before` macro to compare the current `jiffies` value and our timeout.

Or for example if we look into the sound/isa/sscape.c source code file which represents the driver for the Ensoniq Soundscape Elite sound card, we will see the `obp_startup_ack` function that waits upto a given timeout for the On-Board Processor to return its start-up acknowledgement sequence:

```
static int obp_startup_ack(struct soundscape *s, unsigned timeout)
{
    unsigned long end_time = jiffies + msecs_to_jiffies(timeout);

    do {
        ...
        ...
        ...
        x = host_read_unsafe(s->io_base);
        ...
        ...
        ...
        if (x == 0xfe || x == 0xff)
            return 1;
        msleep(10);
    } while (time_before(jiffies, end_time));

    return 0;
}
```

As you can see, the `jiffies` variable is very widely used in the Linux kernel code. As I already wrote, we met yet another new time management related concept in the previous part - `clocksource`. We have only seen a short description of this concept and the API for a clock source registration. Let's take a closer look in this part.

Introduction to `clocksource`

The `clocksource` concept represents the generic API for clock sources management in the Linux kernel. Why do we need a separate framework for this? Let's go back to the beginning. The `time` concept is the fundamental concept in the Linux kernel and other operating system kernels. And the timekeeping is one of the necessities to use this concept. For example Linux kernel must know and update the time elapsed since system startup, it must determine how long the current process has been running for every processor and many many more. Where the Linux kernel can get information about time? First of all it is Real Time Clock or RTC that represents by the a nonvolatile device. You can find a set of architecture-independent real time clock drivers in the Linux kernel in the drivers/rtc directory. Besides this, each architecture can provide a driver for the architecture-dependent real time clock, for example - `CMOS/RTC` - arch/x86/kernel/rtc.c for the x86 architecture. The second is system timer - timer that excites interrupts with a periodic rate. For example, for IBM PC compatibles it was - programmable interval timer.

We already know that for timekeeping purposes we can use `jiffies` in the Linux kernel. The `jiffies` can be considered as read only global variable which is updated with `HZ` frequency. We know that the `HZ` is a compile-time kernel parameter whose reasonable range is from `100` to `1000` Hz. So, it is guaranteed to have an interface for time measurement with `1` - `10` milliseconds resolution. Besides standard `jiffies`, we saw the `refined_jiffies` clock source in the previous part that is based on the `i8253/i8254` programmable interval timer tick rate which is almost `1193182` hertz. So we can get something about `1` microsecond resolution with the `refined_jiffies`. In this time, nanoseconds are the favorite choice for the time value units of the given clock source.

The availability of more precise techniques for time intervals measurement is hardware-dependent. We just knew a little about `x86` dependent timers hardware. But each architecture provides own timers hardware. Earlier each architecture had own implementation for this purpose. Solution of this problem is an abstraction layer and associated API in a common code framework for managing various clock sources and independent of the timer interrupt. This common code framework became - `clocksource` framework.

Generic timeofday and clock source management framework moved a lot of timekeeping code into the architecture independent portion of the code, with the architecture-dependent portion reduced to defining and managing low-level hardware pieces of clocksources. It takes a large amount of funds to measure the time interval on different architectures with different hardware, and it is very complex. Implementation of the each clock related service is strongly associated with an individual hardware device and as you can understand, it results in similar implementations for different architectures.

Within this framework, each clock source is required to maintain a representation of time as a monotonically increasing value. As we can see in the Linux kernel code, nanoseconds are the favorite choice for the time value units of a clock source in this time. One of the main point of the clock source framework is to allow an user to select clock source among a range of available hardware devices supporting clock functions when configuring the system and selecting, accessing and scaling different clock sources.

The clocksource structure

The fundamental of the `clocksource` framework is the `clocksource` structure that defined in the include/linux/clocksource.h header file. We already saw some fields that are provided by the `clocksource` structure in the previous part. Let's look on the full definition of this structure and try to describe all of its fields:

```
struct clocksource {
    cycle_t (*read)(struct clocksource *cs);
    cycle_t mask;
    u32 mult;
    u32 shift;
    u64 max_idle_ns;
    u32 maxadj;
#ifdef CONFIG_ARCH_CLOCKSOURCE_DATA
    struct arch_clocksource_data archdata;
#endif
    u64 max_cycles;
    const char *name;
    struct list_head list;
    int rating;
    int (*enable)(struct clocksource *cs);
    void (*disable)(struct clocksource *cs);
    unsigned long flags;
    void (*suspend)(struct clocksource *cs);
    void (*resume)(struct clocksource *cs);
#ifdef CONFIG_CLOCKSOURCE_WATCHDOG
    struct list_head wd_list;
    cycle_t cs_last;
    cycle_t wd_last;
#endif
    struct module *owner;
} ____cacheline_aligned;
```

We already saw the first field of the `clocksource` structure in the previous part - it is pointer to the `read` function that returns best counter selected by the clocksource framework. For example we use `jiffies_read` function to read `jiffies` value:

```
static struct clocksource clocksource_jiffies = {
    ...
    .read        = jiffies_read,
    ...
}
```

where `jiffies_read` just returns:

```
static cycle_t jiffies_read(struct clocksource *cs)
{
    return (cycle_t) jiffies;
```

```
}
```

Or the `read_tsc` function:

```
static struct clocksource clocksource_tsc = {
    ...
    .read                   = read_tsc,
    ...
};
```

for the time stamp counter reading.

The next field is `mask` that allows to ensure that subtraction between counters values from non `64 bit` counters do not need special overflow logic. After the `mask` field, we can see two fields: `mult` and `shift`. These are the fields that are base of mathematical functions that are provide ability to convert time values specific to each clock source. In other words these two fields help us to convert an abstract machine time units of a counter to nanoseconds.

After these two fields we can see the `64` bits `max_idle_ns` field represents max idle time permitted by the clocksource in nanoseconds. We need in this field for the Linux kernel with enabled `CONFIG_NO_HZ` kernel configuration option. This kernel configuration option enables the Linux kernel to run without a regular timer tick (we will see full explanation of this in other part). The problem that dynamic tick allows the kernel to sleep for periods longer than a single tick, moreover sleep time could be unlimited. The `max_idle_ns` field represents this sleeping limit.

The next field after the `max_idle_ns` is the `maxadj` field which is the maximum adjustment value to `mult`. The main formula by which we convert cycles to the nanoseconds:

```
((u64) cycles * mult) >> shift;
```

is not `100%` accurate. Instead the number is taken as close as possible to a nanosecond and `maxadj` helps to correct this and allows clocksource API to avoid `mult` values that might overflow when adjusted. The next four fields are pointers to the function:

- `enable` - optional function to enable clocksource;
- `disable` - optional function to disable clocksource;
- `suspend` - suspend function for the clocksource;
- `resume` - resume function for the clocksource;

The next field is the `max_cycles` and as we can understand from its name, this field represents maximum cycle value before potential overflow. And the last field is `owner` represents reference to a kernel module that is owner of a clocksource. This is all. We just went through all the standard fields of the `clocksource` structure. But you can noted that we missed some fields of the `clocksource` structure. We can divide all of missed field on two types: Fields of the first type are already known for us. For example, they are `name` field that represents name of a `clocksource`, the `rating` field that helps to the Linux kernel to select the best clocksource and etc. The second type, fields which are dependent from the different Linux kernel configuration options. Let's look on these fields.

The first field is the `archdata`. This field has `arch_clocksource_data` type and depends on the `CONFIG_ARCH_CLOCKSOURCE_DATA` kernel configuration option. This field is actual only for the x86 and IA64 architectures for this moment. And again, as we can understand from the field's name, it represents architecture-specific data for a clock source. For example, it represents `vDSO` clock mode:

```
struct arch_clocksource_data {
    int vclock_mode;
};
```

for the `x86` architectures. Where the `vDSO` clock mode can be one of the:

```
#define VCLOCK_NONE 0
#define VCLOCK_TSC  1
#define VCLOCK_HPET 2
#define VCLOCK_PVCLOCK 3
```

The last three fields are `wd_list` , `cs_last` and the `wd_last` depends on the `CONFIG_CLOCKSOURCE_WATCHDOG` kernel configuration option. First of all let's try to understand what is it `watchdog` . In a simple words, watchdog is a timer that is used for detection of the computer malfunctions and recovering from it. All of these three fields contain watchdog related data that is used by the `clocksource` framework. If we will grep the Linux kernel source code, we will see that only arch/x86/KConfig kernel configuration file contains the `CONFIG_CLOCKSOURCE_WATCHDOG` kernel configuration option. So, why do `x86` and `x86_64` need in watchdog? You already may know that all `x86` processors has special 64-bit register - time stamp counter. This register contains number of cycles since the reset. Sometimes the time stamp counter needs to be verified against another clock source. We will not see initialization of the `watchdog` timer in this part, before this we must learn more about timers.

That's all. From this moment we know all fields of the `clocksource` structure. This knowledge will help us to learn insides of the `clocksource` framework.

New clock source registration

We saw only one function from the `clocksource` framework in the previous part. This function was - `__clocksource_register` . This function defined in the include/linux/clocksource.h header file and as we can understand from the function's name, main point of this function is to register new clocksource. If we will look on the implementation of the `__clocksource_register` function, we will see that it just makes call of the `__clocksource_register_scale` function and returns its result:

```
static inline int __clocksource_register(struct clocksource *cs)
{
    return __clocksource_register_scale(cs, 1, 0);
}
```

Before we will see implementation of the `__clocksource_register_scale` function, we can see that `clocksource` provides additional API for a new clock source registration:

```
static inline int clocksource_register_hz(struct clocksource *cs, u32 hz)
{
        return __clocksource_register_scale(cs, 1, hz);
}

static inline int clocksource_register_khz(struct clocksource *cs, u32 khz)
{
        return __clocksource_register_scale(cs, 1000, khz);
}
```

And all of these functions do the same. They return value of the `__clocksource_register_scale` function but with different set of parameters. The `__clocksource_register_scale` function defined in the kernel/time/clocksource.c source code file. To understand difference between these functions, let's look on the parameters of the `clocksource_register_khz` function. As we can see, this function takes three parameters:

- `cs` - clocksource to be installed;
- `scale` - scale factor of a clock source. In other words, if we will multiply value of this parameter on frequency, we will get `hz` of a clocksource;
- `freq` - clock source frequency divided by scale.

Now let's look on the implementation of the `__clocksource_register_scale` function:

```
int __clocksource_register_scale(struct clocksource *cs, u32 scale, u32 freq)
{
        __clocksource_update_freq_scale(cs, scale, freq);
        mutex_lock(&clocksource_mutex);
        clocksource_enqueue(cs);
        clocksource_enqueue_watchdog(cs);
        clocksource_select();
        mutex_unlock(&clocksource_mutex);
        return 0;
}
```

First of all we can see that the `__clocksource_register_scale` function starts from the call of the `__clocksource_update_freq_scale` function that defined in the same source code file and updates given clock source with the new frequency. Let's look on the implementation of this function. In the first step we need to check given frequency and if it was not passed as `zero`, we need to calculate `mult` and `shift` parameters for the given clock source. Why do we need to check value of the `frequency`? Actually it can be zero. if you attentively looked on the implementation of the `__clocksource_register` function, you may have noticed that we passed `frequency` as `0`. We will do it only for some clock sources that have self defined `mult` and `shift` parameters. Look in the previous part and you will see that we saw calculation of the `mult` and `shift` for `jiffies`. The `__clocksource_update_freq_scale` function will do it for us for other clock sources.

So in the start of the `__clocksource_update_freq_scale` function we check the value of the `frequency` parameter and if is not zero we need to calculate `mult` and `shift` for the given clock source. Let's look on the `mult` and `shift` calculation:

```
void __clocksource_update_freq_scale(struct clocksource *cs, u32 scale, u32 freq)
{
        u64 sec;

        if (freq) {
             sec = cs->mask;
             do_div(sec, freq);
             do_div(sec, scale);

             if (!sec)
                   sec = 1;
             else if (sec > 600 && cs->mask > UINT_MAX)
                   sec = 600;

             clocks_calc_mult_shift(&cs->mult, &cs->shift, freq,
                                    NSEC_PER_SEC / scale, sec * scale);
        }
        ...
        ...
        ...
}
```

Here we can see calculation of the maximum number of seconds which we can run before a clock source counter will overflow. First of all we fill the `sec` variable with the value of a clock source mask. Remember that a clock source's mask represents maximum amount of bits that are valid for the given clock source. After this, we can see two division operations. At first we divide our `sec` variable on a clock source frequency and then on scale factor. The `freq` parameter shows us how many timer interrupts will be occurred in one second. So, we divide `mask` value that represents maximum number of a counter (for example `jiffy`) on the frequency of a timer and will get the maximum number of seconds for the certain clock source. The second division operation will give us maximum number of seconds for the certain clock source depends on its scale factor which can be `1` hertz or `1` kilohertz (10^3 Hz).

After we have got maximum number of seconds, we check this value and set it to `1` or `600` depends on the result at the next step. These values is maximum sleeping time for a clocksource in seconds. In the next step we can see call of the `clocks_calc_mult_shift`. Main point of this function is calculation of the `mult` and `shift` values for a given clock source. In the end of the `__clocksource_update_freq_scale` function we check that just calculated `mult` value of a given clock source will not cause overflow after adjustment, update the `max_idle_ns` and `max_cycles` values of a given clock source with the maximum nanoseconds that can be converted to a clock source counter and print result to the kernel buffer:

```
pr_info("%s: mask: 0x%llx max_cycles: 0x%llx, max_idle_ns: %lld ns\n",
    cs->name, cs->mask, cs->max_cycles, cs->max_idle_ns);
```

that we can see in the dmesg output:

```
$ dmesg | grep "clocksource:"
[    0.000000] clocksource: refined-jiffies: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 191096994039
1419 ns
[    0.000000] clocksource: hpet: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 133484882848 ns
[    0.094084] clocksource: jiffies: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 1911260446275000 ns
[    0.205302] clocksource: acpi_pm: mask: 0xffffff max_cycles: 0xffffff, max_idle_ns: 2085701024 ns
[    1.452979] clocksource: tsc: mask: 0xffffffffffffffff max_cycles: 0x7350b459580, max_idle_ns: 881591204237
ns
```

After the `__clocksource_update_freq_scale` function will finish its work, we can return back to the `__clocksource_register_scale` function that will register new clock source. We can see the call of the following three functions:

```
mutex_lock(&clocksource_mutex);
clocksource_enqueue(cs);
clocksource_enqueue_watchdog(cs);
clocksource_select();
mutex_unlock(&clocksource_mutex);
```

Note that before the first will be called, we lock the `clocksource_mutex` mutex. The point of the `clocksource_mutex` mutex is to protect `curr_clocksource` variable which represents currently selected `clocksource` and `clocksource_list` variable which represents list that contains registered `clocksources` . Now, let's look on these three functions.

The first `clocksource_enqueue` function and other two defined in the same source code file. We go through all already registered `clocksources` or in other words we go through all elements of the `clocksource_list` and tries to find best place for a given `clocksource` :

```
static void clocksource_enqueue(struct clocksource *cs)
{
    struct list_head *entry = &clocksource_list;
    struct clocksource *tmp;

    list_for_each_entry(tmp, &clocksource_list, list)
        if (tmp->rating >= cs->rating)
            entry = &tmp->list;
    list_add(&cs->list, entry);
}
```

In the end we just insert new clocksource to the `clocksource_list` . The second function - `clocksource_enqueue_watchdog` does almost the same that previous function, but it inserts new clock source to the `wd_list` depends on flags of a clock source and starts new watchdog timer. As I already wrote, we will not consider `watchdog` related stuff in this part but will do it in next parts.

The last function is the `clocksource_select` . As we can understand from the function's name, main point of this function - select the best `clocksource` from registered clocksources. This function consists only from the call of the function helper:

```
static void clocksource_select(void)
{
    return __clocksource_select(false);
```

```
}
```

Note that the `__clocksource_select` function takes one parameter (`false` in our case). This bool parameter shows how to traverse the `clocksource_list` . In our case we pass `false` that is meant that we will go through all entries of the `clocksource_list` . We already know that `clocksource` with the best rating will the first in the `clocksource_list` after the call of the `clocksource_enqueue` function, so we can easily get it from this list. After we found a clock source with the best rating, we switch to it:

```
if (curr_clocksource != best && !timekeeping_notify(best)) {
    pr_info("Switched to clocksource %s\n", best->name);
    curr_clocksource = best;
}
```

The result of this operation we can see in the `dmesg` output:

```
$ dmesg | grep Switched
[    0.199688] clocksource: Switched to clocksource hpet
[    2.452966] clocksource: Switched to clocksource tsc
```

Note that we can see two clock sources in the `dmesg` output (`hpet` and `tsc` in our case). Yes, actually there can be many different clock sources on a particular hardware. So the Linux kernel knows about all registered clock sources and switches to a clock source with a better rating each time after registration of a new clock source.

If we will look on the bottom of the kernel/time/clocksource.c source code file, we will see that it has sysfs interface. Main initialization occurs in the `init_clocksource_sysfs` function which will be called during device `initcalls` . Let's look on the implementation of the `init_clocksource_sysfs` function:

```
static struct bus_type clocksource_subsys = {
    .name = "clocksource",
    .dev_name = "clocksource",
};

static int __init init_clocksource_sysfs(void)
{
    int error = subsys_system_register(&clocksource_subsys, NULL);

    if (!error)
        error = device_register(&device_clocksource);
    if (!error)
        error = device_create_file(
                &device_clocksource,
                &dev_attr_current_clocksource);
    if (!error)
        error = device_create_file(&device_clocksource,
                       &dev_attr_unbind_clocksource);
    if (!error)
        error = device_create_file(
                &device_clocksource,
                &dev_attr_available_clocksource);
    return error;
}
device_initcall(init_clocksource_sysfs);
```

First of all we can see that it registers a `clocksource` subsystem with the call of the `subsys_system_register` function. In other words, after the call of this function, we will have following directory:

```
$ pwd
/sys/devices/system/clocksource
```

After this step, we can see registration of the `device_clocksource` device which is represented by the following structure:

```
static struct device device_clocksource = {
    .id    = 0,
    .bus    = &clocksource_subsys,
};
```

and creation of three files:

- `dev_attr_current_clocksource` ;
- `dev_attr_unbind_clocksource` ;
- `dev_attr_available_clocksource` .

These files will provide information about current clock source in the system, available clock sources in the system and interface which allows to unbind the clock source.

After the `init_clocksource_sysfs` function will be executed, we will be able find some information about available clock sources in the:

```
$ cat /sys/devices/system/clocksource/clocksource0/available_clocksource
tsc hpet acpi_pm
```

Or for example information about current clock source in the system:

```
$ cat /sys/devices/system/clocksource/clocksource0/current_clocksource
tsc
```

In the previous part, we saw API for the registration of the `jiffies` clock source, but didn't dive into details about the `clocksource` framework. In this part we did it and saw implementation of the new clock source registration and selection of a clock source with the best rating value in the system. Of course, this is not all API that `clocksource` framework provides. There a couple additional functions like `clocksource_unregister` for removing given clock source from the `clocksource_list` and etc. But I will not describe this functions in this part, because they are not important for us right now. Anyway if you are interesting in it, you can find it in the kernel/time/clocksource.c.

That's all.

Conclusion

This is the end of the second part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part got acquainted with the following two concepts: `jiffies` and `clocksource` . In this part we saw some examples of the `jiffies` usage and knew more details about the `clocksource` concept.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- x86
- x86_64
- uptime
- Ensoniq Soundscape Elite

- RTC
- interrupts
- IBM PC
- programmable interval timer
- Hz
- nanoseconds
- dmesg
- time stamp counter
- loadable kernel module
- IA64
- watchdog
- clock rate
- mutex
- sysfs
- previous part

Timers and time management in the Linux kernel. Part 3.

The tick broadcast framework and dyntick

This is third part of the chapter which describes timers and time management related stuff in the Linux kernel and we stopped on the `clocksource` framework in the previous part. We have started to consider this framework because it is closely related to the special counters which are provided by the Linux kernel. One of these counters which we already saw in the first part of this chapter is - `jiffies`. As I already wrote in the first part of this chapter, we will consider time management related stuff step by step during the Linux kernel initialization. Previous step was call of the:

```
register_refined_jiffies(CLOCK_TICK_RATE);
```

function which defined in the kernel/time/jiffies.c source code file and executes initialization of the `refined_jiffies` clock source for us. Recall that this function is called from the `setup_arch` function that defined in the https://github.com/torvalds/linux/blob/16f73eb02d7e1765ccab3d2018e0bd98eb93d973/arch/x86/kernel/setup.c source code and executes architecture-specific (x86_64 in our case) initialization. Look on the implementation of the `setup_arch` and you will note that the call of the `register_refined_jiffies` is the last step before the `setup_arch` function will finish its work.

There are many different `x86_64` specific things already configured after the end of the `setup_arch` execution. For example some early interrupt handlers already able to handle interrupts, memory space reserved for the initrd, DMI scanned, the Linux kernel log buffer is already set and this means that the printk function is able to work, e820 parsed and the Linux kernel already knows about available memory and and many many other architecture specific things (if you are interesting, you can read more about the `setup_arch` function and Linux kernel initialization process in the second chapter of this book).

Now, the `setup_arch` finished its work and we can back to the generic Linux kernel code. Recall that the `setup_arch` function was called from the `start_kernel` function which is defined in the init/main.c source code file. So, we shall return to this function. You can see that there are many different function are called right after `setup_arch` function inside of the `start_kernel` function, but since our chapter is devoted to timers and time management related stuff, we will skip all code which is not related to this topic. The first function which is related to the time management in the Linux kernel is:

```
tick_init();
```

in the `start_kernel`. The `tick_init` function defined in the kernel/time/tick-common.c source code file and does two things:

- Initialization of `tick broadcast` framework related data structures;
- Initialization of `full` tickless mode related data structures.

We didn't see anything related to the `tick broadcast` framework in this book and didn't know anything about tickless mode in the Linux kernel. So, the main point of this part is to look on these concepts and to know what are they.

The idle process

First of all, let's look on the implementation of the `tick_init` function. As I already wrote, this function defined in the kernel/time/tick-common.c source code file and consists from the two calls of following functions:

```
void __init tick_init(void)
{
    tick_broadcast_init();
    tick_nohz_init();
}
```

As you can understand from the paragraph's title, we are interesting only in the `tick_broadcast_init` function for now. This function defined in the kernel/time/tick-broadcast.c source code file and executes initialization of the `tick broadcast` framework related data structures. Before we will look on the implementation of the `tick_broadcast_init` function and will try to understand what does this function do, we need to know about `tick broadcast` framework.

Main point of a central processor is to execute programs. But sometimes a processor may be in a special state when it is not being used by any program. This special state is called - idle. When the processor has no anything to execute, the Linux kernel launches `idle` task. We already saw a little about this in the last part of the Linux kernel initialization process. When the Linux kernel will finish all initialization processes in the `start_kernel` function from the init/main.c source code file, it will call the `rest_init` function from the same source code file. Main point of this function is to launch kernel `init` thread and the `kthreadd` thread, to call the `schedule` function to start task scheduling and to go to sleep by calling the `cpu_idle_loop` function that defined in the kernel/sched/idle.c source code file.

The `cpu_idle_loop` function represents infinite loop which checks the need for rescheduling on each iteration. After the scheduler finds something to execute, the `idle` process will finish its work and the control will be moved to a new runnable task with the call of the `schedule_preempt_disabled` function:

```
static void cpu_idle_loop(void)
{
    while (1) {
        while (!need_resched()) {
        ...
        ...
        ...
        /* the main idle function */
        cpuidle_idle_call();
    }
    ...
    ...
    ...
    schedule_preempt_disabled();
}
```

Of course, we will not consider full implementation of the `cpu_idle_loop` function and details of the `idle` state in this part, because it is not related to our topic. But there is one interesting moment for us. We know that the processor can execute only one task in one time. How does the Linux kernel decide to reschedule and stop `idle` process if the processor executes infinite loop in the `cpu_idle_loop` ? The answer is system timer interrupts. When an interrupt occurs, the processor stops the `idle` thread and transfers control to an interrupt handler. After the system timer interrupt handler will be handled, the `need_resched` will return true and the Linux kernel will stop `idle` process and will transfer control to the current runnable task. But handling of the system timer interrupts is not effective for power management, because if a processor is in `idle` state, there is little point in sending it a system timer interrupt.

By default, there is the `CONFIG_HZ_PERIODIC` kernel configuration option which is enabled in the Linux kernel and tells to handle each interrupt of the system timer. To solve this problem, the Linux kernel provides two additional ways of managing scheduling-clock interrupts:

The first is to omit scheduling-clock ticks on idle processors. To enable this behaviour in the Linux kernel, we need to enable the `CONFIG_NO_HZ_IDLE` kernel configuration option. This option allows Linux kernel to avoid sending timer interrupts to idle processors. In this case periodic timer interrupts will be replaced with on-demand interrupts. This mode is called - `dyntick-idle` mode. But if the kernel does not handle interrupts of a system timer, how can the kernel decide if the system has nothing to do?

Whenever the idle task is selected to run, the periodic tick is disabled with the call of the `tick_nohz_idle_enter` function that defined in the kernel/time/tick-sched.c source code file and enabled with the call of the `tick_nohz_idle_exit` function. There is special concept in the Linux kernel which is called - `clock event devices` that are used to schedule the next interrupt. This concept provides API for devices which can deliver interrupts at a specific time in the future and represented by the `clock_event_device` structure in the Linux kernel. We will not dive into implementation of the `clock_event_device` structure now. We will see it in the next part of this chapter. But there is one interesting moment for us right now.

The second way is to omit scheduling-clock ticks on processors that are either in `idle` state or that have only one runnable task or in other words busy processor. We can enable this feature with the `CONFIG_NO_HZ_FULL` kernel configuration option and it allows to reduce the number of timer interrupts significantly.

Besides the `cpu_idle_loop`, idle processor can be in a sleeping state. The Linux kernel provides special `cpuidle` framework. Main point of this framework is to put an idle processor to sleeping states. The name of the set of these states is - `C-states`. But how does a processor will be woken if local timer is disabled? The linux kernel provides `tick broadcast` framework for this. The main point of this framework is assign a timer which is not affected by the `C-states`. This timer will wake a sleeping processor.

Now, after some theory we can return to the implementation of our function. Let's recall that the `tick_init` function just calls two following functions:

```
void __init tick_init(void)
{
    tick_broadcast_init();
    tick_nohz_init();
}
```

Let's consider the first function. The first `tick_broadcast_init` function defined in the kernel/time/tick-broadcast.c source code file and executes initialization of the `tick broadcast` framework related data structures. Let's look on the implementation of the `tick_broadcast_init` function:

```
void __init tick_broadcast_init(void)
{
        zalloc_cpumask_var(&tick_broadcast_mask, GFP_NOWAIT);
        zalloc_cpumask_var(&tick_broadcast_on, GFP_NOWAIT);
        zalloc_cpumask_var(&tmpmask, GFP_NOWAIT);
#ifdef CONFIG_TICK_ONESHOT
         zalloc_cpumask_var(&tick_broadcast_oneshot_mask, GFP_NOWAIT);
         zalloc_cpumask_var(&tick_broadcast_pending_mask, GFP_NOWAIT);
         zalloc_cpumask_var(&tick_broadcast_force_mask, GFP_NOWAIT);
#endif
}
```

As we can see, the `tick_broadcast_init` function allocates different cpumasks with the help of the `zalloc_cpumask_var` function. The `zalloc_cpumask_var` function defined in the lib/cpumask.c source code file and expands to the call of the following function:

```
bool zalloc_cpumask_var(cpumask_var_t *mask, gfp_t flags)
{
        return alloc_cpumask_var(mask, flags | __GFP_ZERO);
}
```

Ultimately, the memory space will be allocated for the given `cpumask` with the certain flags with the help of the `kmalloc_node` function:

```
*mask = kmalloc_node(cpumask_size(), flags, node);
```

Now let's look on the `cpumasks` that will be initialized in the `tick_broadcast_init` function. As we can see, the `tick_broadcast_init` function will initialize six `cpumasks`, and moreover, initialization of the last three `cpumasks` will be depended on the `CONFIG_TICK_ONESHOT` kernel configuration option.

The first three `cpumasks` are:

- `tick_broadcast_mask` - the bitmap which represents list of processors that are in a sleeping mode;
- `tick_broadcast_on` - the bitmap that stores numbers of processors which are in a periodic broadcast state;

- `tmpmask` - this bitmap for temporary usage.

As we already know, the next three `cpumasks` depends on the `CONFIG_TICK_ONESHOT` kernel configuration option. Actually each clock event devices can be in one of two modes:

- `periodic` - clock events devices that support periodic events;
- `oneshot` - clock events devices that capable of issuing events that happen only once.

The linux kernel defines two mask for such clock events devices in the include/linux/clockchips.h header file:

```
#define CLOCK_EVT_FEAT_PERIODIC        0x000001
#define CLOCK_EVT_FEAT_ONESHOT         0x000002
```

So, the last three `cpumasks` are:

- `tick_broadcast_oneshot_mask` - stores numbers of processors that must be notified;
- `tick_broadcast_pending_mask` - stores numbers of processors that pending broadcast;
- `tick_broadcast_force_mask` - stores numbers of processors with enforced broadcast.

We have initialized six `cpumasks` in the `tick broadcast` framework, and now we can proceed to implementation of this framework.

The `tick broadcast` framework

Hardware may provide some clock source devices. When a processor sleeps and its local timer stopped, there must be additional clock source device that will handle awakening of a processor. The Linux kernel uses these `special` clock source devices which can raise an interrupt at a specified time. We already know that such timers called `clock events` devices in the Linux kernel. Besides `clock events` devices, each processor in the system has its own local timer which is programmed to issue interrupt at the time of the next deferred task. Also these timers can be programmed to do a periodical job, like updating `jiffies` and etc. These timers represented by the `tick_device` structure in the Linux kernel. This structure defined in the kernel/time/tick-sched.h header file and looks:

```
struct tick_device {
        struct clock_event_device *evtdev;
        enum tick_device_mode mode;
};
```

Note, that the `tick_device` structure contains two fields. The first field - `evtdev` represents pointer to the `clock_event_device` structure that defined in the include/linux/clockchips.h header file and represents descriptor of a clock event device. A `clock event` device allows to register an event that will happen in the future. As I already wrote, we will not consider `clock_event_device` structure and related API in this part, but will see it in the next part.

The second field of the `tick_device` structure represents mode of the `tick_device`. As we already know, the mode can be one of the:

```
enum tick_device_mode {
        TICKDEV_MODE_PERIODIC,
        TICKDEV_MODE_ONESHOT,
};
```

Each `clock events` device in the system registers itself by the call of the `clockevents_register_device` function or `clockevents_config_and_register` function during initialization process of the Linux kernel. During the registration of a new `clock events` device, the Linux kernel calls the `tick_check_new_device` function that defined in the kernel/time/tick-common.c source code file and checks the given `clock events` device should be used by the Linux kernel. After all checks, the `tick_check_new_device` function executes a call of the:

```
tick_install_broadcast_device(newdev);
```

function that checks that the given `clock event` device can be broadcast device and install it, if the given device can be broadcast device. Let's look on the implementation of the `tick_install_broadcast_device` function:

```
void tick_install_broadcast_device(struct clock_event_device *dev)
{
    struct clock_event_device *cur = tick_broadcast_device.evtdev;

    if (!tick_check_broadcast_device(cur, dev))
        return;

    if (!try_module_get(dev->owner))
        return;

    clockevents_exchange_device(cur, dev);

    if (cur)
        cur->event_handler = clockevents_handle_noop;

    tick_broadcast_device.evtdev = dev;

    if (!cpumask_empty(tick_broadcast_mask))
        tick_broadcast_start_periodic(dev);

    if (dev->features & CLOCK_EVT_FEAT_ONESHOT)
        tick_clock_notify();
}
```

First of all we get the current `clock event` device from the `tick_broadcast_device`. The `tick_broadcast_device` defined in the kernel/time/tick-common.c source code file:

```
static struct tick_device tick_broadcast_device;
```

and represents external clock device that keeps track of events for a processor. The first step after we got the current clock device is the call of the `tick_check_broadcast_device` function which checks that a given clock events device can be utilized as broadcast device. The main point of the `tick_check_broadcast_device` function is to check value of the `features` field of the given `clock events` device. As we can understand from the name of this field, the `features` field contains a clock event device features. Available values defined in the include/linux/clockchips.h header file and can be one of the `CLOCK_EVT_FEAT_PERIODIC` - which represents a clock events device which supports periodic events and etc. So, the `tick_check_broadcast_device` function check `features` flags for `CLOCK_EVT_FEAT_ONESHOT`, `CLOCK_EVT_FEAT_DUMMY` and other flags and returns `false` if the given clock events device has one of these features. In other way the `tick_check_broadcast_device` function compares `ratings` of the given clock event device and current clock event device and returns the best.

After the `tick_check_broadcast_device` function, we can see the call of the `try_module_get` function that checks module owner of the clock events. We need to do it to be sure that the given `clock events` device was correctly initialized. The next step is the call of the `clockevents_exchange_device` function that defined in the kernel/time/clockevents.c source code file and will release old clock events device and replace the previous functional handler with a dummy handler.

In the last step of the `tick_install_broadcast_device` function we check that the `tick_broadcast_mask` is not empty and start the given `clock events` device in periodic mode with the call of the `tick_broadcast_start_periodic` function:

```
if (!cpumask_empty(tick_broadcast_mask))
    tick_broadcast_start_periodic(dev);

if (dev->features & CLOCK_EVT_FEAT_ONESHOT)
    tick_clock_notify();
```

The `tick_broadcast_mask` filled in the `tick_device_uses_broadcast` function that checks a `clock events` device during registration of this `clock events` device:

```
int cpu = smp_processor_id();

int tick_device_uses_broadcast(struct clock_event_device *dev, int cpu)
{
    ...
    ...
    ...
    if (!tick_device_is_functional(dev)) {
        ...
        cpumask_set_cpu(cpu, tick_broadcast_mask);
        ...
    }
    ...
    ...
    ...
}
```

More about the `smp_processor_id` macro you can read in the fourth part of the Linux kernel initialization process chapter.

The `tick_broadcast_start_periodic` function check the given `clock event` device and call the `tick_setup_periodic` function:

```
static void tick_broadcast_start_periodic(struct clock_event_device *bc)
{
    if (bc)
        tick_setup_periodic(bc, 1);
}
```

that defined in the kernel/time/tick-common.c source code file and sets broadcast handler for the given `clock event` device by the call of the following function:

```
tick_set_periodic_handler(dev, broadcast);
```

This function checks the second parameter which represents broadcast state (`on` or `off`) and sets the broadcast handler depends on its value:

```
void tick_set_periodic_handler(struct clock_event_device *dev, int broadcast)
{
    if (!broadcast)
        dev->event_handler = tick_handle_periodic;
    else
        dev->event_handler = tick_handle_periodic_broadcast;
}
```

When an `clock event` device will issue an interrupt, the `dev->event_handler` will be called. For example, let's look on the interrupt handler of the high precision event timer which is located in the arch/x86/kernel/hpet.c source code file:

```
static irqreturn_t hpet_interrupt_handler(int irq, void *data)
{
    struct hpet_dev *dev = (struct hpet_dev *)data;
    struct clock_event_device *hevt = &dev->evt;

    if (!hevt->event_handler) {
        printk(KERN_INFO "Spurious HPET timer interrupt on HPET timer %d\n",
                dev->num);
        return IRQ_HANDLED;
```

```
    }

    hevt->event_handler(hevt);
    return IRQ_HANDLED;
}
```

The `hpet_interrupt_handler` gets the irq specific data and check the event handler of the `clock event` device. Recall that we just set in the `tick_set_periodic_handler` function. So the `tick_handler_periodic_broadcast` function will be called in the end of the high precision event timer interrupt handler.

The `tick_handler_periodic_broadcast` function calls the

```
bc_local = tick_do_periodic_broadcast();
```

function which stores numbers of processors which have asked to be woken up in the temporary `cpumask` and call the `tick_do_broadcast` function:

```
cpumask_and(tmpmask, cpu_online_mask, tick_broadcast_mask);
return tick_do_broadcast(tmpmask);
```

The `tick_do_broadcast` calls the `broadcast` function of the given clock events which sends IPI interrupt to the set of the processors. In the end we can call the event handler of the given `tick_device` :

```
if (bc_local)
    td->evtdev->event_handler(td->evtdev);
```

which actually represents interrupt handler of the local timer of a processor. After this a processor will wake up. That is all about `tick broadcast` framework in the Linux kernel. We have missed some aspects of this framework, for example reprogramming of a `clock event` device and broadcast with the oneshot timer and etc. But the Linux kernel is very big, it is not real to cover all aspects of it. I think it will be interesting to dive into with yourself.

If you remember, we have started this part with the call of the `tick_init` function. We just consider the `tick_broadcast_init` function and related theory, but the `tick_init` function contains another call of a function and this function is - `tick_nohz_init` . Let's look on the implementation of this function.

Initialization of dyntick related data structures

We already saw some information about `dyntick` concept in this part and we know that this concept allows kernel to disable system timer interrupts in the `idle` state. The `tick_nohz_init` function makes initialization of the different data structures which are related to this concept. This function defined in the kernel/time/tick-sched.c source code file and starts from the check of the value of the `tick_nohz_full_running` variable which represents state of the tick-less mode for the `idle` state and the state when system timer interrupts are disabled during a processor has only one runnable task:

```
if (!tick_nohz_full_running) {
    if (tick_nohz_init_all() < 0)
    return;
}
```

If this mode is not running we call the `tick_nohz_init_all` function that defined in the same source code file and check its result. The `tick_nohz_init_all` function tries to allocate the `tick_nohz_full_mask` with the call of the `alloc_cpumask_var` that will allocate space for a `tick_nohz_full_mask` . The `tick_nohz_full_mask` will store numbers of processors that have enabled full `NO_HZ` . After successful allocation of the `tick_nohz_full_mask` we set all bits in the `tick_nohz_full_mask` , set the `tick_nohz_full_running` and return result to the `tick_nohz_init` function:

```C
static int tick_nohz_init_all(void)
{
        int err = -1;
#ifdef CONFIG_NO_HZ_FULL_ALL
        if (!alloc_cpumask_var(&tick_nohz_full_mask, GFP_KERNEL)) {
                WARN(1, "NO_HZ: Can't allocate full dynticks cpumask\n");
                return err;
        }
        err = 0;
        cpumask_setall(tick_nohz_full_mask);
        tick_nohz_full_running = true;
#endif
        return err;
}
```

In the next step we try to allocate a memory space for the `housekeeping_mask` :

```C
if (!alloc_cpumask_var(&housekeeping_mask, GFP_KERNEL)) {
    WARN(1, "NO_HZ: Can't allocate not-full dynticks cpumask\n");
    cpumask_clear(tick_nohz_full_mask);
    tick_nohz_full_running = false;
    return;
}
```

This `cpumask` will store number of processor for `housekeeping` or in other words we need at least in one processor that will not be in `NO_HZ` mode, because it will do timekeeping and etc. After this we check the result of the architecture-specific `arch_irq_work_has_interrupt` function. This function checks ability to send inter-processor interrupt for the certain architecture. We need to check this, because system timer of a processor will be disabled during `NO_HZ` mode, so there must be at least one online processor which can send inter-processor interrupt to awake offline processor. This function defined in the arch/x86/include/asm/irq_work.h header file for the x86_64 and just checks that a processor has APIC from the CPUID:

```C
static inline bool arch_irq_work_has_interrupt(void)
{
    return cpu_has_apic;
}
```

If a processor has not `APIC` , the Linux kernel prints warning message, clears the `tick_nohz_full_mask` cpumask, copies numbers of all possible processors in the system to the `housekeeping_mask` and resets the value of the `tick_nohz_full_running` variable:

```C
if (!arch_irq_work_has_interrupt()) {
    pr_warning("NO_HZ: Can't run full dynticks because arch doesn't "
           "support irq work self-IPIs\n");
    cpumask_clear(tick_nohz_full_mask);
    cpumask_copy(housekeeping_mask, cpu_possible_mask);
    tick_nohz_full_running = false;
    return;
}
```

After this step, we get the number of the current processor by the call of the `smp_processor_id` and check this processor in the `tick_nohz_full_mask` . If the `tick_nohz_full_mask` contains a given processor we clear appropriate bit in the `tick_nohz_full_mask` :

```C
cpu = smp_processor_id();

if (cpumask_test_cpu(cpu, tick_nohz_full_mask)) {
    pr_warning("NO_HZ: Clearing %d from nohz_full range for timekeeping\n", cpu);
    cpumask_clear_cpu(cpu, tick_nohz_full_mask);
}
```

Because this processor will be used for timekeeping. After this step we put all numbers of processors that are in the `cpu_possible_mask` and not in the `tick_nohz_full_mask` :

```
cpumask_andnot(housekeeping_mask,
           cpu_possible_mask, tick_nohz_full_mask);
```

After this operation, the `housekeeping_mask` will contain all processors of the system except a processor for timekeeping. In the last step of the `tick_nohz_init_all` function, we are going through all processors that are defined in the `tick_nohz_full_mask` and call the following function for an each processor:

```
for_each_cpu(cpu, tick_nohz_full_mask)
    context_tracking_cpu_set(cpu);
```

The `context_tracking_cpu_set` function defined in the kernel/context_tracking.c source code file and main point of this function is to set the `context_tracking.active` percpu variable to `true` . When the `active` field will be set to `true` for the certain processor, all context switches will be ignored by the Linux kernel context tracking subsystem for this processor.

That's all. This is the end of the `tick_nohz_init` function. After this `NO_HZ` related data structures will be initialized. We didn't see API of the `NO_HZ` mode, but will see it soon.

Conclusion

This is the end of the third part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part got acquainted with the `clocksource` concept in the Linux kernel which represents framework for managing different clock source in a interrupt and hardware characteristics independent way. We continued to look on the Linux kernel initialization process in a time management context in this part and got acquainted with two new concepts for us: the `tick broadcast` framework and `tick-less` mode. The first concept helps the Linux kernel to deal with processors which are in deep sleep and the second concept represents the mode in which kernel may work to improve power management of `idle` processors.

In the next part we will continue to dive into timer management related things in the Linux kernel and will see new concept for us - `timers` .

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- x86_64
- initrd
- interrupt
- DMI
- printk
- CPU idle
- power management
- NO_HZ documentation
- cpumasks
- high precision event timer
- irq
- IPI

* CPUID
* APIC
* percpu
* context switches
* Previous part

Timers and time management in the Linux kernel. Part 4.

Timers

This is fourth part of the chapter which describes timers and time management related stuff in the Linux kernel and in the previous part we knew about the `tick broadcast` framework and `NO_HZ` mode in the Linux kernel. We will continue to dive into the time management related stuff in the Linux kernel in this part and will be acquainted with yet another concept in the Linux kernel - `timers` . Before we will look at timers in the Linux kernel, we have to learn some theory about this concept. Note that we will consider software timers in this part.

The Linux kernel provides a `software timer` concept to allow to kernel functions could be invoked at future moment. Timers are widely used in the Linux kernel. For example, look in the net/netfilter/ipset/ip_set_list_set.c source code file. This source code file provides implementation of the framework for the managing of groups of IP addresses.

We can find the `list_set` structure that contains `gc` filed in this source code file:

```
struct list_set {
    ...
    struct timer_list gc;
    ...
};
```

Not that the `gc` filed has `timer_list` type. This structure defined in the include/linux/timer.h header file and main point of this structure is to store `dynamic` timers in the Linux kernel. Actually, the Linux kernel provides two types of timers called dynamic timers and interval timers. First type of timers is used by the kernel, and the second can be used by user mode. The `timer_list` structure contains actual `dynamic` timers. The `list_set` contains `gc` timer in our example represents timer for garbage collection. This timer will be initialized in the `list_set_gc_init` function:

```
static void
list_set_gc_init(struct ip_set *set, void (*gc)(unsigned long ul_set))
{
    struct list_set *map = set->data;
    ...
    ...
    ...
    map->gc.function = gc;
    map->gc.expires = jiffies + IPSET_GC_PERIOD(set->timeout) * HZ;
    ...
    ...
    ...
}
```

A function that is pointed by the `gc` pointer, will be called after timeout which is equal to the `map->gc.expires` .

Ok, we will not dive into this example with the netfilter, because this chapter is not about network related stuff. But we saw that timers are widely used in the Linux kernel and learned that they represent concept which allows to functions to be called in future.

Now let's continue to research source code of Linux kernel which is related to the timers and time management stuff as we did it in all previous chapters.

Introduction to dynamic timers in the Linux kernel

As I already wrote, we knew about the `tick broadcast` framework and `NO_HZ` mode in the previous part. They will be initialized in the init/main.c source code file by the call of the `tick_init` function. If we will look at this source code file, we will see that the next time management related function is:

```
init_timers();
```

This function defined in the kernel/time/timer.c source code file and contains calls of four functions:

```
void __init init_timers(void)
{
    init_timer_cpus();
    init_timer_stats();
    timer_register_cpu_notifier();
    open_softirq(TIMER_SOFTIRQ, run_timer_softirq);
}
```

Let's look on implementation of each function. The first function is `init_timer_cpus` defined in the same source code file and just calls the `init_timer_cpu` function for each possible processor in the system:

```
static void __init init_timer_cpus(void)
{
    int cpu;

    for_each_possible_cpu(cpu)
        init_timer_cpu(cpu);
}
```

If you do not know or do not remember what is it a `possible` cpu, you can read the special part of this book which describes `cpumask` concept in the Linux kernel. In short words, a `possible` processor is a processor which can be plugged in anytime during the life of the system.

The `init_timer_cpu` function does main work for us, namely it executes initialization of the `tvec_base` structure for each processor. This structure defined in the kernel/time/timer.c source code file and stores data related to a `dynamic` timer for a certain processor. Let's look on the definition of this structure:

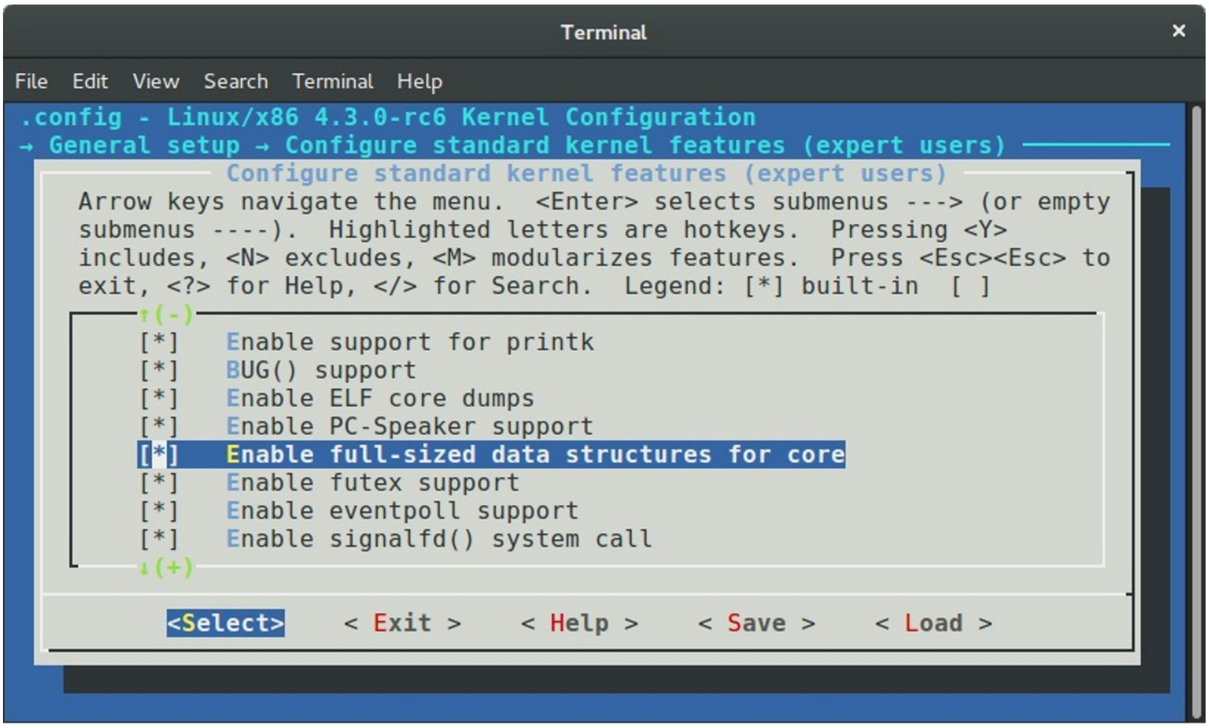
```
struct tvec_base {
    spinlock_t lock;
    struct timer_list *running_timer;
    unsigned long timer_jiffies;
    unsigned long next_timer;
    unsigned long active_timers;
    unsigned long all_timers;
    int cpu;
    bool migration_enabled;
    bool nohz_active;
    struct tvec_root tv1;
    struct tvec tv2;
    struct tvec tv3;
    struct tvec tv4;
    struct tvec tv5;
} ____cacheline_aligned;
```

The `thec_base` structure contains following fields: The `lock` for `tvec_base` protection, the next `running_timer` field points to the currently running timer for the certain processor, the `timer_jiffies` fields represents the earliest expiration time (it will be used by the Linux kernel to find already expired timers). The next field - `next_timer` contains the next pending timer for a next timer interrupt in a case when a processor goes to sleep and the `NO_HZ` mode is enabled in the Linux kernel. The `active_timers` field provides accounting of non-deferrable timers or in other words all timers that will not be stopped during a processor will go to sleep. The `all_timers` field tracks total number of timers or `active_timers` + deferrable timers. The `cpu` field represents number of a processor which owns timers. The `migration_enabled` and `nohz_active` fields are represent opportunity of timers migration to another processor and status of the `NO_HZ` mode respectively.

The last five fields of the `tvec_base` structure represent lists of dynamic timers. The first `tv1` field has:

```
#define TVR_SIZE (1 << TVR_BITS)
#define TVR_BITS (CONFIG_BASE_SMALL ? 6 : 8)

...
...
...

struct tvec_root {
    struct hlist_head vec[TVR_SIZE];
};
```

type. Note that the value of the `TVR_SIZE` depends on the `CONFIG_BASE_SMALL` kernel configuration option:



that reduces size of the kernel data structures if disabled. The `v1` is array that may contain `64` or `256` elements where an each element represents a dynamic timer that will decay within the next `255` system timer interrupts. Next three fields: `tv2` , `tv3` and `tv4` are lists with dynamic timers too, but they store dynamic timers which will decay the next `2^14 - 1` , `2^20 - 1` and `2^26` respectively. The last `tv5` field represents list which stores dynamic timers with a large expiring period.

So, now we saw the `tvec_base` structure and description of its fields and we can look on the implementation of the `init_timer_cpu` function. As I already wrote, this function defined in the kernel/time/timer.c source code file and executes initialization of the `tvec_bases` :

```
static void __init init_timer_cpu(int cpu)
{
    struct tvec_base *base = per_cpu_ptr(&tvec_bases, cpu);

    base->cpu = cpu;
    spin_lock_init(&base->lock);

    base->timer_jiffies = jiffies;
    base->next_timer = base->timer_jiffies;
```

```
}
```

The `tvec_bases` represents per-cpu variable which represents main data structure for a dynamic timer for a given processor. This `per-cpu` variable defined in the same source code file:

```
static DEFINE_PER_CPU(struct tvec_base, tvec_bases);
```

First of all we're getting the address of the `tvec_bases` for the given processor to `base` variable and as we got it, we are starting to initialize some of the `tvec_base` fields in the `init_timer_cpu` function. After initialization of the `per-cpu` dynamic timers with the jiffies and the number of a possible processor, we need to initialize a `tstats_lookup_lock` spinlock in the `init_timer_stats` function:

```
void __init init_timer_stats(void)
{
    int cpu;

    for_each_possible_cpu(cpu)
        raw_spin_lock_init(&per_cpu(tstats_lookup_lock, cpu));
}
```

The `tstats_lookcup_lock` variable represents `per-cpu` raw spinlock:

```
static DEFINE_PER_CPU(raw_spinlock_t, tstats_lookup_lock);
```

which will be used for protection of operation with statistics of timers that can be accessed through the procfs:

```
static int __init init_tstats_procfs(void)
{
    struct proc_dir_entry *pe;

    pe = proc_create("timer_stats", 0644, NULL, &tstats_fops);
    if (!pe)
        return -ENOMEM;
    return 0;
}
```

For example:

```
$ cat /proc/timer_stats
Timerstats sample period: 3.888770 s
  12,     0 swapper          hrtimer_stop_sched_tick (hrtimer_sched_tick)
  15,     1 swapper          hcd_submit_urb (rh_timer_func)
   4,   959 kedac            schedule_timeout (process_timeout)
   1,     0 swapper          page_writeback_init (wb_timer_fn)
  28,     0 swapper          hrtimer_stop_sched_tick (hrtimer_sched_tick)
  22,  2948 IRQ 4            tty_flip_buffer_push (delayed_work_timer_fn)
  ...
  ...
  ...
```

The next step after initialization of the `tstats_lookup_lock` spinlock is the call of the `timer_register_cpu_notifier` function. This function depends on the `CONFIG_HOTPLUG_CPU` kernel configuration option which enables support for hotplug processors in the Linux kernel.

When a processor will be logically offlined, a notification will be sent to the Linux kernel with the `CPU_DEAD` or the `CPU_DEAD_FROZEN` event by the call of the `cpu_notifier` macro:

```
#ifdef CONFIG_HOTPLUG_CPU
...
...
static inline void timer_register_cpu_notifier(void)
{
    cpu_notifier(timer_cpu_notify, 0);
}
...
...
#else
...
...
static inline void timer_register_cpu_notifier(void) { }
...
...
#endif /* CONFIG_HOTPLUG_CPU */
```

In this case the `timer_cpu_notify` will be called which checks an event type and will call the `migrate_timers` function:

```
static int timer_cpu_notify(struct notifier_block *self,
                            unsigned long action, void *hcpu)
{
    switch (action) {
    case CPU_DEAD:
    case CPU_DEAD_FROZEN:
        migrate_timers((long)hcpu);
        break;
    default:
        break;
    }

    return NOTIFY_OK;
}
```

This chapter will not describe `hotplug` related events in the Linux kernel source code, but if you are interesting in such things, you can find implementation of the `migrate_timers` function in the kernel/time/timer.c source code file.

The last step in the `init_timers` function is the call of the:

```
open_softirq(TIMER_SOFTIRQ, run_timer_softirq);
```

function. The `open_softirq` function may be already familiar to you if you have read the ninth part about the interrupts and interrupt handling in the Linux kernel. In short words, the `open_softirq` function defined in the kernel/softirq.c source code file and executes initialization of the deferred interrupt handler.

In our case the deferred function is the `run_timer_softirq` function that is will be called after a hardware interrupt in the `do_IRQ` function which defined in the arch/x86/kernel/irq.c source code file. The main point of this function is to handle a software dynamic timer. The Linux kernel does not do this thing during the hardware timer interrupt handling because this is time consuming operation.

Let's look on the implementation of the `run_timer_softirq` function:

```
static void run_timer_softirq(struct softirq_action *h)
{
    struct tvec_base *base = this_cpu_ptr(&tvec_bases);

    if (time_after_eq(jiffies, base->timer_jiffies))
        __run_timers(base);
}
```

At the beginning of the `run_timer_softirq` function we get a `dynamic` timer for a current processor and compares the current value of the jiffies with the value of the `timer_jiffies` for the current structure by the call of the `time_after_eq` macro which is defined in the include/linux/jiffies.h header file:

```
#define time_after_eq(a,b)          \
    (typecheck(unsigned long, a) && \
     typecheck(unsigned long, b) && \
    ((long)((a) - (b)) >= 0))
```

Reclaim that the `timer_jiffies` field of the `tvec_base` structure represents the relative time when functions delayed by the given timer will be executed. So we compare these two values and if the current time represented by the `jiffies` is greater than `base->timer_jiffies`, we call the `__run_timers` function that defined in the same source code file. Let's look on the implementation of this function.

As I just wrote, the `__run_timers` function runs all expired timers for a given processor. This function starts from the acquiring of the `tvec_base's` lock to protect the `tvec_base` structure

```
static inline void __run_timers(struct tvec_base *base)
{
    struct timer_list *timer;

    spin_lock_irq(&base->lock);
    ...
    ...
    ...
    spin_unlock_irq(&base->lock);
}
```

After this it starts the loop while the `timer_jiffies` will not be greater than the `jiffies`:

```
while (time_after_eq(jiffies, base->timer_jiffies)) {
    ...
    ...
    ...
}
```

We can find many different manipulations in the our loop, but the main point is to find expired timers and call delayed functions. First of all we need to calculate the `index` of the `base->tv1` list that stores the next timer to be handled with the following expression:

```
index = base->timer_jiffies & TVR_MASK;
```

where the `TVR_MASK` is a mask for the getting of the `tvec_root->vec` elements. As we got the index with the next timer which must be handled we check its value. If the index is zero, we go through all lists in our cascade table `tv2`, `tv3` and etc., and rehashing it with the call of the `cascade` function:

```
if (!index &&
    (!cascade(base, &base->tv2, INDEX(0))) &&
        (!cascade(base, &base->tv3, INDEX(1))) &&
                !cascade(base, &base->tv4, INDEX(2)))
        cascade(base, &base->tv5, INDEX(3));
```

After this we increase the value of the `base->timer_jiffies`:

```
++base->timer_jiffies;
```

In the last step we are executing a corresponding function for each timer from the list in a following loop:

```
hlist_move_list(base->tv1.vec + index, head);

while (!hlist_empty(head)) {
    ...
    ...
    ...
    timer = hlist_entry(head->first, struct timer_list, entry);
    fn = timer->function;
    data = timer->data;

    spin_unlock(&base->lock);
    call_timer_fn(timer, fn, data);
    spin_lock(&base->lock);

    ...
    ...
    ...
}
```

where the `call_timer_fn` just call the given function:

```
static void call_timer_fn(struct timer_list *timer, void (*fn)(unsigned long),
                          unsigned long data)
{
    ...
    ...
    ...
    fn(data);
    ...
    ...
    ...
}
```

That's all. The Linux kernel has infrastructure for `dynamic timers` from this moment. We will not dive into this interesting theme. As I already wrote the `timers` is a widely used concept in the Linux kernel and nor one part, nor two parts will not cover understanding of such things how it implemented and how it works. But now we know about this concept, why does the Linux kernel needs in it and some data structures around it.

Now let's look usage of `dynamic timers` in the Linux kernel.

Usage of dynamic timers

As you already can noted, if the Linux kernel provides a concept, it also provides API for managing of this concept and the `dynamic timers` concept is not exception here. To use a timer in the Linux kernel code, we must define a variable with a `timer_list` type. We can initialize our `timer_list` structure in two ways. The first is to use the `init_timer` macro that defined in the include/linux/timer.h header file:

```
#define init_timer(timer)    \
    __init_timer((timer), 0)

#define __init_timer(_timer, _flags)   \
         init_timer_key((_timer), (_flags), NULL, NULL)
```

where the `init_timer_key` function just calls the:

```
do_init_timer(timer, flags, name, key);
```

function which fields the given `timer` with default values. The second way is to use the:

```
#define TIMER_INITIALIZER(_function, _expires, _data)        \
    __TIMER_INITIALIZER((_function), (_expires), (_data), 0)
```

macro which will initialize the given `timer_list` structure too.

After a `dynamic timer` is initialized we can start this `timer` with the call of the:

```
void add_timer(struct timer_list * timer);
```

function and stop it with the:

```
int del_timer(struct timer_list * timer);
```

function.

That's all.

Conclusion

This is the end of the fourth part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part we got acquainted with the two new concepts: the `tick broadcast` framework and the `NO_HZ` mode. In this part we continued to dive into time management related stuff and got acquainted with the new concept - `dynamic timer` or software timer. We didn't saw implementation of a `dynamic timers` management code in details in this part but saw data structures and API around this concept.

In the next part we will continue to dive into timer management related things in the Linux kernel and will see new concept for us - `timers`.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- IP
- netfilter
- network
- cpumask
- interrupt
- jiffies
- per-cpu
- spinlock
- procfs
- previous part

Timers and time management in the Linux kernel. Part 5.

Introduction to the `clockevents` framework

This is fifth part of the chapter which describes timers and time management related stuff in the Linux kernel. As you might noted from the title of this part, the `clockevents` framework will be discussed. We already saw one framework in the second part of this chapter. It was `clocksource` framework. Both of these frameworks represent timekeeping abstractions in the Linux kernel.

At first let's refresh your memory and try to remember what is it `clocksource` framework and and what its purpose. The main goal of the `clocksource` framework is to provide `timeline`. As described in the documentation:

> For example issuing the command 'date' on a Linux system will eventually read the clock source to determine exactly what time it is.

The Linux kernel supports many different clock sources. You can find some of them in the drivers/clocksource. For example old good Intel 8253 - programmable interval timer with `1193182` Hz frequency, yet another one - ACPI PM timer with `3579545` Hz frequency. Besides the drivers/clocksource directory, each architecture may provide own architecture-specific clock sources. For example x86 architecture provides High Precision Event Timer, or for example powerpc provides access to the processor timer through `timebase` register.

Each clock source provides monotonic atomic counter. As I already wrote, the Linux kernel supports a huge set of different clock source and each clock source has own parameters like frequency. The main goal of the `clocksource` framework is to provide API to select best available clock source in the system i.e. a clock source with the highest frequency. Additional goal of the `clocksource` framework is to represent an atomic counter provided by a clock source in human units. In this time, nanoseconds are the favorite choice for the time value units of the given clock source in the Linux kernel.

The `clocksource` framework represented by the `clocksource` structure which is defined in the include/linux/clocksource.h header code file which contains `name` of a clock source, rating of certain clock source in the system (a clock source with the higher frequency has the biggest rating in the system), `list` of all registered clock source in the system, `enable` and `disable` fields to enable and disable a clock source, pointer to the `read` function which must return an atomic counter of a clock source and etc.

Additionally the `clocksource` structure provides two fields: `mult` and `shift` which are needed for translation of an atomic counter which is provided by a certain clock source to the human units, i.e. nanoseconds. Translation occurs via following formula:

```
ns ~= (clocksource * mult) >> shift
```

As we already know, besides the `clocksource` structure, the `clocksource` framework provides an API for registration of clock source with different frequency scale factor:

```
static inline int clocksource_register_hz(struct clocksource *cs, u32 hz)
static inline int clocksource_register_khz(struct clocksource *cs, u32 khz)
```

A clock source unregistration:

```
int clocksource_unregister(struct clocksource *cs)
```

and etc.

Additionally to the `clocksource` framework, the Linux kernel provides `clockevents` framework. As described in the documentation:

> Clock events are the conceptual reverse of clock sources

Main goal of the is to manage clock event devices or in other words - to manage devices that allow to register an event or in other words interrupt that is going to happen at a defined point of time in the future.

Now we know a little about the `clockevents` framework in the Linux kernel, and now time is to see on it API.

API of `clockevents` framework

The main structure which described a clock event device is `clock_event_device` structure. This structure is defined in the include/linux/clockchips.h header file and contains a huge set of fields. as well as the `clocksource` structure it has `name` fields which contains human readable name of a clock event device, for example local APIC timer:

```
static struct clock_event_device lapic_clockevent = {
    .name                   = "lapic",
    ...
    ...
    ...
}
```

Addresses of the `event_handler` , `set_next_event` , `next_event` functions for a certain clock event device which are an interrupt handler, setter of next event and local storage for next event respectively. Yet another field of the `clock_event_device` structure is - `features` field. Its value maybe on of the following generic features:

```
#define CLOCK_EVT_FEAT_PERIODIC    0x000001
#define CLOCK_EVT_FEAT_ONESHOT        0x000002
```

Where the `CLOCK_EVT_FEAT_PERIODIC` represents device which may be programmed to generate events periodically. The `CLOCK_EVT_FEAT_ONESHOT` represents device which may generate an event only once. Besides these two features, there are also architecture-specific features. For example x86_64 supports two additional features:

```
#define CLOCK_EVT_FEAT_C3STOP        0x000008
```

The first `CLOCK_EVT_FEAT_C3STOP` means that a clock event device will be stopped in the C3 state. Additionally the `clock_event_device` structure has `mult` and `shift` fields as well as `clocksource` structure. The `clocksource` structure also contains other fields, but we will consider it later.

After we considered part of the `clock_event_device` structure, time is to look at the `API` of the `clockevents` framework. To work with a clock event device, first of all we need to initialize `clock_event_device` structure and register a clock events device. The `clockevents` framework provides following `API` for registration of clock event devices:

```
void clockevents_register_device(struct clock_event_device *dev)
{
   ...
   ...
   ...
}
```

This function defined in the kernel/time/clockevents.c source code file and as we may see, the `clockevents_register_device` function takes only one parameter:

- address of a `clock_event_device` structure which represents a clock event device.

So, to register a clock event device, at first we need to initialize `clock_event_device` structure with parameters of a certain clock event device. Let's take a look at one random clock event device in the Linux kernel source code. We can find one in the drivers/clocksource directory or try to take a look at an architecture-specific clock event device. Let's take for example - Periodic Interval Timer (PIT) for at91sam926x. You can find its implementation in the drivers/clocksource.

First of all let's look at initialization of the `clock_event_device` structure. This occurs in the `at91sam926x_pit_common_init` function:

```
struct pit_data {
    ...
    ...
    struct clock_event_device       clkevt;
    ...
    ...
};

static void __init at91sam926x_pit_common_init(struct pit_data *data)
{
    ...
    ...
    ...
    data->clkevt.name = "pit";
    data->clkevt.features = CLOCK_EVT_FEAT_PERIODIC;
    data->clkevt.shift = 32;
    data->clkevt.mult = div_sc(pit_rate, NSEC_PER_SEC, data->clkevt.shift);
    data->clkevt.rating = 100;
    data->clkevt.cpumask = cpumask_of(0);

    data->clkevt.set_state_shutdown = pit_clkevt_shutdown;
    data->clkevt.set_state_periodic = pit_clkevt_set_periodic;
    data->clkevt.resume = at91sam926x_pit_resume;
    data->clkevt.suspend = at91sam926x_pit_suspend;
    ...
}
```

Here we can see that `at91sam926x_pit_common_init` takes one parameter - pointer to the `pit_data` structure which contains `clock_event_device` structure which will contain clock event related information of the `at91sam926x` periodic Interval Timer. At the start we fill `name` of the timer device and its `features`. In our case we deal with periodic timer which as we already know may be programmed to generate events periodically.

The next two fields `shift` and `mult` are familiar to us. They will be used to translate counter of our timer to nanoseconds. After this we set rating of the timer to `100`. This means if there will not be timers with higher rating in the system, this timer will be used for timekeeping. The next field - `cpumask` indicates for which processors in the system the device will work. In our case, the device will work for the first processor. The `cpumask_of` macro defined in the include/linux/cpumask.h header file and just expands to the call of the:

```
#define cpumask_of(cpu) (get_cpu_mask(cpu))
```

Where the `get_cpu_mask` returns the cpumask containing just a given `cpu` number. More about `cpumasks` concept you may read in the CPU masks in the Linux kernel part. In the last four lines of code we set callbacks for the clock event device suspend/resume, device shutdown and update of the clock event device state.

After we finished with the initialization of the `at91sam926x` periodic timer, we can register it by the call of the following functions:

```
clockevents_register_device(&data->clkevt);
```

Now we can consider implementation of the `clockevent_register_device` function. As I already wrote above, this function is defined in the kernel/time/clockevents.c source code file and starts from the initialization of the initial event device state:

```
clockevent_set_state(dev, CLOCK_EVT_STATE_DETACHED);
```

Actually, an event device may be in one of this states:

```
enum clock_event_state {
    CLOCK_EVT_STATE_DETACHED,
    CLOCK_EVT_STATE_SHUTDOWN,
    CLOCK_EVT_STATE_PERIODIC,
    CLOCK_EVT_STATE_ONESHOT,
    CLOCK_EVT_STATE_ONESHOT_STOPPED,
};
```

Where:

- `CLOCK_EVT_STATE_DETACHED` - a clock event device is not not used by `clockevents` framework. Actually it is initial state of all clock event devices;
- `CLOCK_EVT_STATE_SHUTDOWN` - a clock event device is powered-off;
- `CLOCK_EVT_STATE_PERIODIC` - a clock event device may be programmed to generate event periodically;
- `CLOCK_EVT_STATE_ONESHOT` - a clock event device may be programmed to generate event only once;
- `CLOCK_EVT_STATE_ONESHOT_STOPPED` - a clock event device was programmed to generate event only once and now it is temporary stopped.

The implementation of the `clock_event_set_state` function is pretty easy:

```
static inline void clockevent_set_state(struct clock_event_device *dev,
                    enum clock_event_state state)
{
    dev->state_use_accessors = state;
}
```

As we can see, it just fills the `state_use_accessors` field of the given `clock_event_device` structure with the given value which is in our case is `CLOCK_EVT_STATE_DETACHED`. Actually all clock event devices has this initial state during registration. The `state_use_accessors` field of the `clock_event_device` structure provides `current` state of the clock event device.

After we have set initial state of the given `clock_event_device` structure we check that the `cpumask` of the given clock event device is not zero:

```
if (!dev->cpumask) {
    WARN_ON(num_possible_cpus() > 1);
    dev->cpumask = cpumask_of(smp_processor_id());
}
```

Remember that we have set the `cpumask` of the `at91sam926x` periodic timer to first processor. If the `cpumask` field is zero, we check the number of possible processors in the system and print warning message if it is less than on. Additionally we set the `cpumask` of the given clock event device to the current processor. If you are interested in how the `smp_processor_id` macro is implemented, you can read more about it in the fourth part of the Linux kernel initialization process chapter.

After this check we lock the actual code of the clock event device registration by the call following macros:

```
raw_spin_lock_irqsave(&clockevents_lock, flags);
...
...
...
raw_spin_unlock_irqrestore(&clockevents_lock, flags);
```

Additionally the `raw_spin_lock_irqsave` and the `raw_spin_unlock_irqrestore` macros disable local interrupts, however interrupts on other processors still may occur. We need to do it to prevent potential deadlock if we adding new clock event device to the list of clock event devices and an interrupt occurs from other clock event device.

We can see following code of clock event device registration between the `raw_spin_lock_irqsave` and `raw_spin_unlock_irqrestore` macros:

```
list_add(&dev->list, &clockevent_devices);
tick_check_new_device(dev);
clockevents_notify_released();
```

First of all we add the given clock event device to the list of clock event devices which is represented by the `clockevent_devices` :

```
static LIST_HEAD(clockevent_devices);
```

At the next step we call the `tick_check_new_device` function which is defined in the kernel/time/tick-common.c source code file and checks do the new registered clock event device should be used or not. The `tick_check_new_device` function checks the given `clock_event_device` gets the current registered tick device which is represented by the `tick_device` structure and compares their ratings and features. Actually `CLOCK_EVT_STATE_ONESHOT` is preferred:

```
static bool tick_check_preferred(struct clock_event_device *curdev,
                 struct clock_event_device *newdev)
{
    if (!(newdev->features & CLOCK_EVT_FEAT_ONESHOT)) {
        if (curdev && (curdev->features & CLOCK_EVT_FEAT_ONESHOT))
            return false;
        if (tick_oneshot_mode_active())
            return false;
    }

    return !curdev ||
        newdev->rating > curdev->rating ||
           !cpumask_equal(curdev->cpumask, newdev->cpumask);
}
```

If the new registered clock event device is more preferred than old tick device, we exchange old and new registered devices and install new device:

```
clockevents_exchange_device(curdev, newdev);
tick_setup_device(td, newdev, cpu, cpumask_of(cpu));
```

The `clockevents_exchange_device` function releases or in other words deleted the old clock event device from the `clockevent_devices` list. The next function - `tick_setup_device` as we may understand from its name, setups new tick device. This function check the mode of the new registered clock event device and call the `tick_setup_periodic` function or the `tick_setup_oneshot` depends on the tick device mode:

```
if (td->mode == TICKDEV_MODE_PERIODIC)
    tick_setup_periodic(newdev, 0);
else
    tick_setup_oneshot(newdev, handler, next_event);
```

Both of this functions calls the `clockevents_switch_state` to change state of the clock event device and the `clockevents_program_event` function to set next event of clock event device based on delta between the maximum and minimum difference current time and time for the next event. The `tick_setup_periodic` :

```
clockevents_switch_state(dev, CLOCK_EVT_STATE_PERIODIC);
clockevents_program_event(dev, next, false))
```

and the `tick_setup_oneshot_periodic` :

```
clockevents_switch_state(newdev, CLOCK_EVT_STATE_ONESHOT);
clockevents_program_event(newdev, next_event, true);
```

The `clockevents_switch_state` function checks that the clock event device is not in the given state and calls the `__clockevents_switch_state` function from the same source code file:

```
if (clockevent_get_state(dev) != state) {
    if (__clockevents_switch_state(dev, state))
        return;
```

The `__clockevents_switch_state` function just makes a call of the certain callback depends on the given state:

```
static int __clockevents_switch_state(struct clock_event_device *dev,
                      enum clock_event_state state)
{
    if (dev->features & CLOCK_EVT_FEAT_DUMMY)
        return 0;

    switch (state) {
    case CLOCK_EVT_STATE_DETACHED:
    case CLOCK_EVT_STATE_SHUTDOWN:
        if (dev->set_state_shutdown)
            return dev->set_state_shutdown(dev);
        return 0;

    case CLOCK_EVT_STATE_PERIODIC:
        if (!(dev->features & CLOCK_EVT_FEAT_PERIODIC))
            return -ENOSYS;
        if (dev->set_state_periodic)
            return dev->set_state_periodic(dev);
        return 0;
    ...
    ...
    ...
```

In our case for `at91sam926x` periodic timer, the state is the `CLOCK_EVT_FEAT_PERIODIC` :

```
data->clkevt.features = CLOCK_EVT_FEAT_PERIODIC;
data->clkevt.set_state_periodic = pit_clkevt_set_periodic;
```

So, for the `pit_clkevt_set_periodic` callback will be called. If we will read the documentation of the Periodic Interval Timer (PIT) for at91sam926x, we will see that there is `Periodic Interval Timer Mode Register` which allows us to control of periodic interval timer.

It looks like:

```
31                                                    25        24
+---------------------------------------------------------------+
|                                          |  PITIEN  |  PITEN  |
+---------------------------------------------------------------+
23                             19                             16
+---------------------------------------------------------------+
|                             |               PIV              |
+---------------------------------------------------------------+
```

```
15                                                              8
+--------------------------------------------------------------+
|                            PIV                               |
+--------------------------------------------------------------+
7                                                              0
+--------------------------------------------------------------+
|                            PIV                               |
+--------------------------------------------------------------+
```

Where `PIV` or `Periodic Interval Value` - defines the value compared with the primary `20-bit` counter of the Periodic Interval Timer. The `PITEN` or `Period Interval Timer Enabled` if the bit is `1` and the `PITIEN` or `Periodic Interval Timer Interrupt Enable` if the bit is `1`. So, to set periodic mode, we need to set `24`, `25` bits in the `Periodic Interval Timer Mode Register`. And we are doing it in the `pit_clkevt_set_periodic` function:

```
static int pit_clkevt_set_periodic(struct clock_event_device *dev)
{
        struct pit_data *data = clkevt_to_pit_data(dev);
        ...
        ...
        ...
        pit_write(data->base, AT91_PIT_MR,
                  (data->cycle - 1) | AT91_PIT_PITEN | AT91_PIT_PITIEN);

        return 0;
}
```

Where the `AT91_PT_MR`, `AT91_PT_PITEN` and the `AT91_PIT_PITIEN` are declared as:

```
#define AT91_PIT_MR             0x00
#define AT91_PIT_PITIEN       BIT(25)
#define AT91_PIT_PITEN        BIT(24)
```

After the setup of the new clock event device is finished, we can return to the `clockevents_register_device` function. The last function in the `clockevents_register_device` function is:

```
clockevents_notify_released();
```

This function checks the `clockevents_released` list which contains released clock event devices (remember that they may occur after the call of the `clockevents_exchange_device` function). If this list is not empty, we go through clock event devices from the `clock_events_released` list and delete it from the `clockevent_devices`:

```
static void clockevents_notify_released(void)
{
    struct clock_event_device *dev;

    while (!list_empty(&clockevents_released)) {
        dev = list_entry(clockevents_released.next,
                 struct clock_event_device, list);
        list_del(&dev->list);
        list_add(&dev->list, &clockevent_devices);
        tick_check_new_device(dev);
    }
}
```

That's all. From this moment we have registered new clock event device. So the usage of the `clockevents` framework is simple and clear. Architectures registered their clock event devices, in the clock events core. Users of the clockevents core can get clock event devices for their use. The `clockevents` framework provides notification mechanisms for various clock related management events like a clock event device registered or unregistered, a processor is offlined in system which supports CPU hotplug and etc.

We saw implementation only of the `clockevents_register_device` function. But generally, the clock event layer API is small. Besides the `API` for clock event device registration, the `clockevents` framework provides functions to schedule the next event interrupt, clock event device notification service and support for suspend and resume for clock event devices.

If you want to know more about `clockevents` API you can start to research following source code and header files: kernel/time/tick-common.c, kernel/time/clockevents.c and include/linux/clockchips.h.

That's all.

Conclusion

This is the end of the fifth part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part got acquainted with the `timers` concept. In this part we continued to learn time management related stuff in the Linux kernel and saw a little about yet another framework - `clockevents`.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- timekeeping documentation
- Intel 8253
- programmable interval timer
- ACPI pdf
- x86
- High Precision Event Timer
- powerpc
- frequency
- API
- nanoseconds
- interrupt
- interrupt handler
- local APIC
- C3 state
- Periodic Interval Timer (PIT) for at91sam926x
- CPU masks in the Linux kernel
- deadlock
- CPU hotplug
- previous part

Timers and time management in the Linux kernel. Part 6.

x86_64 related clock sources

This is sixth part of the chapter which describes timers and time management related stuff in the Linux kernel. In the previous part we saw `clockevents` framework and now we will continue to dive into time management related stuff in the Linux kernel. This part will describe implementation of x86 architecture related clock sources (more about `clocksource` concept you can read in the second part of this chapter).'

First of all we must know what clock sources may be used at `x86` architecture. It is easy to know from the sysfs or from content of the `/sys/devices/system/clocksource/clocksource0/available_clocksource`. The `/sys/devices/system/clocksource/clocksourceN` provides two special files to achieve this:

- `available_clocksource` - provides information about available clock sources in the system;
- `current_clocksource` - provides information about currently used clock source in the system.

So, let's look:

```
$ cat /sys/devices/system/clocksource/clocksource0/available_clocksource
tsc hpet acpi_pm
```

We can see that there are three registered clock sources in my system:

- `tsc` - Time Stamp Counter;
- `hpet` - High Precision Event Timer;
- `acpi_pm` - ACPI Power Management Timer.

Now let's look at the second file which provides best clock source (a clock source which has the best rating in the system):

```
$ cat /sys/devices/system/clocksource/clocksource0/current_clocksource
tsc
```

For me it is Time Stamp Counter. As we may know from the second part of this chapter, which describes internals of the `clocksource` framework in the Linux kernel, the best clock source in a system is a clock source with the best (highest) rating or in other words with the highest frequency.

Frequency of the ACPI power management timer is `3.579545 MHz`. Frequency of the High Precision Event Timer is at least `10 MHz`. And the frequency of the Time Stamp Counter depends on processor. For example On older processors, the `Time Stamp Counter` was counting internal processor clock cycles. This means its frequency changed when the processor's frequency scaling changed. The situation has changed for newer processors. Newer processors have an `invariant Time Stamp counter` that increments at a constant rate in all operational states of processor. Actually we can get its frequency in the output of the `/proc/cpuinfo`. For example for the first processor in the system:

```
$ cat /proc/cpuinfo
...
model name    : Intel(R) Core(TM) i7-4790K CPU @ 4.00GHz
...
```

And although Intel manual says that the frequency of the `Time Stamp Counter`, while constant, is not necessarily the maximum qualified frequency of the processor, or the frequency given in the brand string, anyway we may see that it will be much more than frequency of the `ACPI PM` timer or `High Precision Event Timer`. And we can see that the clock source with the best rating or highest frequency is current in the system.

You can note that besides these three clock source, we don't see yet another two familiar us clock sources in the output of the `/sys/devices/system/clocksource/clocksource0/available_clocksource`. These clock sources are `jiffy` and `refined_jiffies`. We don't see them because this filed maps only high resolution clock sources or in other words clock sources with the CLOCK_SOURCE_VALID_FOR_HRES flag.

As I already wrote above, we will consider all of these three clock sources in this part. We will consider it in order of their initialization or:

- `hpet`;
- `acpi_pm`;
- `tsc`.

We can make sure that the order is exactly like this in the output of the dmesg util:

```
$ dmesg | grep clocksource
[    0.000000] clocksource: refined-jiffies: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 191096994039
1419 ns
[    0.000000] clocksource: hpet: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 133484882848 ns
[    0.094369] clocksource: jiffies: mask: 0xffffffff max_cycles: 0xffffffff, max_idle_ns: 1911260446275000 ns
[    0.186498] clocksource: Switched to clocksource hpet
[    0.196827] clocksource: acpi_pm: mask: 0xffffff max_cycles: 0xffffff, max_idle_ns: 2085701024 ns
[    1.413685] tsc: Refined TSC clocksource calibration: 3999.981 MHz
[    1.413688] clocksource: tsc: mask: 0xffffffffffffffff max_cycles: 0x73509721780, max_idle_ns: 881591102108
ns
[    2.413748] clocksource: Switched to clocksource tsc
```

The first clock source is the High Precision Event Timer, so let's start from it.

High Precision Event Timer

The implementation of the `High Precision Event Timer` for the x86 architecture is located in the arch/x86/kernel/hpet.c source code file. Its initialization starts from the call of the `hpet_enable` function. This function is called during Linux kernel initialization. If we will look into `start_kernel` function from the init/main.c source code file, we will see that after the all architecture-specific stuff initialized, early console is disabled and time management subsystem already ready, call of the following function:

```
if (late_time_init)
    late_time_init();
```

which does initialization of the late architecture specific timers after early jiffy counter already initialized. The definition of the `late_time_init` function for the `x86` architecture is located in the arch/x86/kernel/time.c source code file. It looks pretty easy:

```
static __init void x86_late_time_init(void)
{
    x86_init.timers.timer_init();
    tsc_init();
}
```

As we may see, it does initialization of the `x86` related timer and initialization of the `Time Stamp Counter`. The seconds we will see in the next paragraph, but now let's consider the call of the `x86_init.timers.timer_init` function. The `timer_init` points to the `hpet_time_init` function from the same source code file. We can verify this by looking on the definition of the `x86_init` structure from the arch/x86/kernel/x86_init.c:

```
struct x86_init_ops x86_init __initdata = {
   ...
   ...
```

```
    ...
    .timers = {
        .setup_percpu_clockev    = setup_boot_APIC_clock,
        .timer_init        = hpet_time_init,
        .wallclock_init        = x86_init_noop,
    },
    ...
    ...
    ...
```

The `hpet_time_init` function does setup of the programmable interval timer if we can not enable `High Precision Event Timer` and setups default timer IRQ for the enabled timer:

```
void __init hpet_time_init(void)
{
    if (!hpet_enable())
        setup_pit_timer();
    setup_default_timer_irq();
}
```

First of all the `hpet_enable` function check we can enable `High Precision Event Timer` in the system by the call of the `is_hpet_capable` function and if we can, we map a virtual address space for it:

```
int __init hpet_enable(void)
{
    if (!is_hpet_capable())
        return 0;

    hpet_set_mapping();
}
```

The `is_hpet_capable` function checks that we didn't pass `hpet=disable` to the kernel command line and the `hpet_address` is received from the ACPI HPET table. The `hpet_set_mapping` function just maps the virtual address spaces for the timer registers:

```
hpet_virt_address = ioremap_nocache(hpet_address, HPET_MMAP_SIZE);
```

As we can read in the IA-PC HPET (High Precision Event Timers) Specification:

> The timer register space is 1024 bytes

So, the `HPET_MMAP_SIZE` is `1024` bytes too:

```
#define HPET_MMAP_SIZE        1024
```

After we mapped virtual space for the `High Precision Event Timer`, we read `HPET_ID` register to get number of the timers:

```
id = hpet_readl(HPET_ID);

last = (id & HPET_ID_NUMBER) >> HPET_ID_NUMBER_SHIFT;
```

We need to get this number to allocate correct amount of space for the `General Configuration Register` of the `High Precision Event Timer`:

```
cfg = hpet_readl(HPET_CFG);

hpet_boot_cfg = kmalloc((last + 2) * sizeof(*hpet_boot_cfg), GFP_KERNEL);
```

After the space is allocated for the configuration register of the `High Precision Event Timer`, we allow to main counter to run, and allow timer interrupts if they are enabled by the setting of `HPET_CFG_ENABLE` bit in the configuration register for all timers. In the end we just register new clock source by the call of the `hpet_clocksource_register` function:

```
if (hpet_clocksource_register())
    goto out_nohpet;
```

which just calls already familiar

```
clocksource_register_hz(&clocksource_hpet, (u32)hpet_freq);
```

function. Where the `clocksource_hpet` is the `clocksource` structure with the rating `250` (remember rating of the previous `refined_jiffies` clock source was `2`), name - `hpet` and `read_hpet` callback for the reading of atomic counter provided by the `High Precision Event Timer`:

```
static struct clocksource clocksource_hpet = {
    .name        = "hpet",
    .rating        = 250,
    .read        = read_hpet,
    .mask        = HPET_MASK,
    .flags        = CLOCK_SOURCE_IS_CONTINUOUS,
    .resume        = hpet_resume_counter,
    .archdata    = { .vclock_mode = VCLOCK_HPET },
};
```

After the `clocksource_hpet` is registered, we can return to the `hpet_time_init()` function from the arch/x86/kernel/time.c source code file. We can remember that the last step is the call of the:

```
setup_default_timer_irq();
```

function in the `hpet_time_init()`. The `setup_default_timer_irq` function checks existence of `legacy` IRQs or in other words support for the i8259 and setups IRQ0 depends on this.

That's all. From this moment the High Precision Event Timer clock source registered in the Linux kernel `clock source` framework and may be used from generic kernel code via the `read_hpet`:

```
static cycle_t read_hpet(struct clocksource *cs)
{
    return (cycle_t)hpet_readl(HPET_COUNTER);
}
```

function which just reads and returns atomic counter from the `Main Counter Register`.

ACPI PM timer

The seconds clock source is ACPI Power Management Timer. Implementation of this clock source is located in the drivers/clocksource/acpi_pm.c source code file and starts from the call of the `init_acpi_pm_clocksource` function during `fs` initcall.

If we will look at implementation of the `init_acpi_pm_clocksource` function, we will see that it starts from the check of the value of `pmtmr_ioport` variable:

```
static int __init init_acpi_pm_clocksource(void)
{
```

```
    ...
    ...
    ...
    if (!pmtmr_ioport)
        return -ENODEV;
    ...
    ...
    ...
```

This `pmtmr_ioport` variable contains extended address of the `Power Management Timer Control Register Block`. It gets its value in the `acpi_parse_fadt` function which is defined in the arch/x86/kernel/acpi/boot.c source code file. This function parses `FADT` or `Fixed ACPI Description Table` ACPI table and tries to get the values of the `X_PM_TMR_BLK` field which contains extended address of the `Power Management Timer Control Register Block`, represented in `Generic Address Structure` format:

```
static int __init acpi_parse_fadt(struct acpi_table_header *table)
{
#ifdef CONFIG_X86_PM_TIMER
        ...
        ...
        ...
        pmtmr_ioport = acpi_gbl_FADT.xpm_timer_block.address;
        ...
        ...
        ...
#endif
    return 0;
}
```

So, if the `CONFIG_X86_PM_TIMER` Linux kernel configuration option is disabled or something going wrong in the `acpi_parse_fadt` function, we can't access the `Power Management Timer` register and return from the `init_acpi_pm_clocksource`. In other way, if the value of the `pmtmr_ioport` variable is not zero, we check rate of this timer and register this clock source by the call of the:

```
clocksource_register_hz(&clocksource_acpi_pm, PMTMR_TICKS_PER_SEC);
```

function. After the call of the `clocksource_register_hs`, the `acpi_pm` clock source will be registered in the `clocksource` framework of the Linux kernel:

```
static struct clocksource clocksource_acpi_pm = {
    .name        = "acpi_pm",
    .rating        = 200,
    .read        = acpi_pm_read,
    .mask        = (cycle_t)ACPI_PM_MASK,
    .flags        = CLOCK_SOURCE_IS_CONTINUOUS,
};
```

with the rating - `200` and the `acpi_pm_read` callback to read atomic counter provided by the `acpi_pm` clock source. The `acpi_pm_read` function just executes `read_pmtmr` function:

```
static cycle_t acpi_pm_read(struct clocksource *cs)
{
    return (cycle_t)read_pmtmr();
}
```

which reads value of the `Power Management Timer` register. This register has following structure:

```
+-------------------------------+----------------------------------+
```

```
|                               |                                  |
|  upper eight bits of a        |      running count of the        |
| 32-bit power management timer |     power management timer       |
|                               |                                  |
+-------------------------------+----------------------------------+
31          E_TMR_VAL           24               TMR_VAL           0
```

Address of this register is stored in the `Fixed ACPI Description Table` ACPI table and we already have it in the `pmtmr_ioport`. So, the implementation of the `read_pmtmr` function is pretty easy:

```
static inline u32 read_pmtmr(void)
{
    return inl(pmtmr_ioport) & ACPI_PM_MASK;
}
```

We just read the value of the `Power Management Timer` register and mask its `24` bits.

That's all. Now we move to the last clock source in this part - `Time Stamp Counter`.

Time Stamp Counter

The third and last clock source in this part is - Time Stamp Counter clock source and its implementation is located in the arch/x86/kernel/tsc.c source code file. We already saw the `x86_late_time_init` function in this part and initialization of the Time Stamp Counter starts from this place. This function calls the `tsc_init()` function from the arch/x86/kernel/tsc.c source code file.

At the beginning of the `tsc_init` function we can see check, which checks that a processor has support of the `Time Stamp Counter`:

```
void __init tsc_init(void)
{
    u64 lpj;
    int cpu;

    if (!cpu_has_tsc) {
        setup_clear_cpu_cap(X86_FEATURE_TSC_DEADLINE_TIMER);
        return;
    }
    ...
    ...
    ...
```

The `cpu_has_tsc` macro expands to the call of the `cpu_has` macro:

```
#define cpu_has_tsc        boot_cpu_has(X86_FEATURE_TSC)

#define boot_cpu_has(bit)    cpu_has(&boot_cpu_data, bit)

#define cpu_has(c, bit)                            \
    (__builtin_constant_p(bit) && REQUIRED_MASK_BIT_SET(bit) ? 1 :    \
     test_cpu_cap(c, bit))
```

which check the given bit (the `X86_FEATURE_TSC_DEADLINE_TIMER` in our case) in the `boot_cpu_data` array which is filled during early Linux kernel initialization. If the processor has support of the `Time Stamp Counter`, we get the frequency of the `Time Stamp Counter` by the call of the `calibrate_tsc` function from the same source code file which tries to get frequency from the different source like Model Specific Register, calibrate over programmable interval timer and etc, after this we initialize frequency and scale factor for the all processors in the system:

```
tsc_khz = x86_platform.calibrate_tsc();
cpu_khz = tsc_khz;

for_each_possible_cpu(cpu) {
    cyc2ns_init(cpu);
    set_cyc2ns_scale(cpu_khz, cpu);
}
```

because only first bootstrap processor will call the `tsc_init`. After this we check hat `Time Stamp Counter` is not disabled:

```
if (tsc_disabled > 0)
    return;
...
...
...
check_system_tsc_reliable();
```

and call the `check_system_tsc_reliable` function which sets the `tsc_clocksource_reliable` if bootstrap processor has the `X86_FEATURE_TSC_RELIABLE` feature. Note that we went through the `tsc_init` function, but did not register our clock source. Actual registration of the `Time Stamp Counter` clock source occurs in the:

```
static int __init init_tsc_clocksource(void)
{
    if (!cpu_has_tsc || tsc_disabled > 0 || !tsc_khz)
        return 0;
    ...
    ...
    ...
    if (boot_cpu_has(X86_FEATURE_TSC_RELIABLE)) {
        clocksource_register_khz(&clocksource_tsc, tsc_khz);
        return 0;
    }
```

function. This function called during the `device` initcall. We do it to be sure that the `Time Stamp Counter` clock source will be registered after the High Precision Event Timer clock source.

After these all three clock sources will be registered in the `clocksource` framework and the `Time Stamp Counter` clock source will be selected as active, because it has the highest rating among other clock sources:

```
static struct clocksource clocksource_tsc = {
    .name                   = "tsc",
    .rating                 = 300,
    .read                   = read_tsc,
    .mask                   = CLOCKSOURCE_MASK(64),
    .flags                  = CLOCK_SOURCE_IS_CONTINUOUS | CLOCK_SOURCE_MUST_VERIFY,
    .archdata               = { .vclock_mode = VCLOCK_TSC },
};
```

That's all.

Conclusion

This is the end of the sixth part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part got acquainted with the `clockevents` framework. In this part we continued to learn time management related stuff in the Linux kernel and saw a little about three different clock sources which are used in the x86 architecture. The next part will be last part of this chapter and we will see some user space related stuff, i.e. how some time related system calls implemented in the Linux kernel.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- x86
- sysfs
- Time Stamp Counter
- High Precision Event Timer
- ACPI Power Management Timer (PDF)
- frequency.
- dmesg
- programmable interval timer
- IRQ
- IA-PC HPET (High Precision Event Timers) Specification
- IRQ0
- i8259
- initcall
- previous part

Timers and time management in the Linux kernel. Part 7.

Time related system calls in the Linux kernel

This is the seventh and last part chapter, which describes timers and time management related stuff in the Linux kernel. In the previous part, we discussed timers in the context of x86_64: High Precision Event Timer and Time Stamp Counter. Internal time management is an interesting part of the Linux kernel, but of course not only the kernel needs the `time` concept. Our programs also need to know time. In this part, we will consider implementation of some time management related system calls. These system calls are:

- `clock_gettime` ;
- `gettimeofday` ;
- `nanosleep` .

We will start from a simple userspace C program and see all way from the call of the standard library function to the implementation of certain system calls. As each architecture provides its own implementation of certain system calls, we will consider only x86_64 specific implementations of system calls, as this book is related to this architecture.

Additionally, we will not consider the concept of system calls in this part, but only implementations of these three system calls in the Linux kernel. If you are interested in what is a `system call` , there is a special chapter about this.

So, let's start from the `gettimeofday` system call.

Implementation of the `gettimeofday` system call

As we can understand from the name `gettimeofday` , this function returns the current time. First of all, let's look at the following simple example:

```
#include <time.h>
#include <sys/time.h>
#include <stdio.h>

int main(int argc, char **argv)
{
    char buffer[40];
    struct timeval time;

    gettimeofday(&time, NULL);

    strftime(buffer, 40, "Current date/time: %m-%d-%Y/%T", localtime(&time.tv_sec));
    printf("%s\n",buffer);

    return 0;
}
```

As you can see, here we call the `gettimeofday` function, which takes two parameters. The first parameter is a pointer to the `timeval` structure, which represents an elapsed time:

```
struct timeval {
    time_t      tv_sec;     /* seconds */
    suseconds_t tv_usec;    /* microseconds */
};
```

The second parameter of the `gettimeofday` function is a pointer to the `timezone` structure which represents a timezone. In our example, we pass address of the `timeval time` to the `gettimeofday` function, the Linux kernel fills the given `timeval` structure and returns it back to us. Additionally, we format the time with the `strftime` function to get something more human readable than elapsed microseconds. Let's see the result:

```
~$ gcc date.c -o date
~$ ./date
Current date/time: 03-26-2016/16:42:02
```

As you may already know, a userspace application does not call a system call directly from the kernel space. Before the actual system call entry will be called, we call a function from the standard library. In my case it is glibc, so I will consider this case. The implementation of the `gettimeofday` function is located in the sysdeps/unix/sysv/linux/x86/gettimeofday.c source code file. As you already may know, the `gettimeofday` is not a usual system call. It is located in the special area which is called `vDSO` (you can read more about it in the part, which describes this concept).

The `glibc` implementation of `gettimeofday` tries to resolve the given symbol; in our case this symbol is `__vdso_gettimeofday` by the call of the `_dl_vdso_vsym` internal function. If the symbol cannot be resolved, it returns `NULL` and we fallback to the call of the usual system call:

```
return (_dl_vdso_vsym ("__vdso_gettimeofday", &linux26)
  ?: (void*) (&__gettimeofday_syscall));
```

The `gettimeofday` entry is located in the arch/x86/entry/vdso/vclock_gettime.c source code file. As we can see the `gettimeofday` is a weak alias of the `__vdso_gettimeofday` :

```
int gettimeofday(struct timeval *, struct timezone *)
    __attribute__((weak, alias("__vdso_gettimeofday")));
```

The `__vdso_gettimeofday` is defined in the same source code file and calls the `do_realtime` function if the given `timeval` is not null:

```
notrace int __vdso_gettimeofday(struct timeval *tv, struct timezone *tz)
{
    if (likely(tv != NULL)) {
        if (unlikely(do_realtime((struct timespec *)tv) == VCLOCK_NONE))
            return vdso_fallback_gtod(tv, tz);
        tv->tv_usec /= 1000;
    }
    if (unlikely(tz != NULL)) {
        tz->tz_minuteswest = gtod->tz_minuteswest;
        tz->tz_dsttime = gtod->tz_dsttime;
    }

    return 0;
}
```

If the `do_realtime` will fail, we fallback to the real system call via call the `syscall` instruction and passing the `__NR_gettimeofday` system call number and the given `timeval` and `timezone` :

```
notrace static long vdso_fallback_gtod(struct timeval *tv, struct timezone *tz)
{
    long ret;

    asm("syscall" : "=a" (ret) :
        "0" (__NR_gettimeofday), "D" (tv), "S" (tz) : "memory");
    return ret;
}
```

The `do_realtime` function gets the time data from the `vsyscall_gtod_data` structure which is defined in the arch/x86/include/asm/vgtod.h header file and contains mapping of the `timespec` structure and a couple of fields which are related to the current clock source in the system. This function fills the given `timeval` structure with values from the `vsyscall_gtod_data` which contains a time related data which is updated via timer interrupt.

First of all we try to access the `gtod` or `global time of day` the `vsyscall_gtod_data` structure via the call of the `gtod_read_begin` and will continue to do it until it will be successful:

```
do {
    seq = gtod_read_begin(gtod);
    mode = gtod->vclock_mode;
    ts->tv_sec = gtod->wall_time_sec;
    ns = gtod->wall_time_snsec;
    ns += vgetsns(&mode);
    ns >>= gtod->shift;
} while (unlikely(gtod_read_retry(gtod, seq)));

ts->tv_sec += __iter_div_u64_rem(ns, NSEC_PER_SEC, &ns);
ts->tv_nsec = ns;
```

As we got access to the `gtod`, we fill the `ts->tv_sec` with the `gtod->wall_time_sec` which stores current time in seconds gotten from the real time clock during initialization of the timekeeping subsystem in the Linux kernel and the same value but in nanoseconds. In the end of this code we just fill the given `timespec` structure with the resulted values.

That's all about the `gettimeofday` system call. The next system call in our list is the `clock_gettime`.

Implementation of the clock_gettime system call

The `clock_gettime` function gets the time which is specified by the second parameter. Generally the `clock_gettime` function takes two parameters:

- `clk_id` - clock identifier;
- `timespec` - address of the `timespec` structure which represent elapsed time.

Let's look on the following simple example:

```
#include <time.h>
#include <sys/time.h>
#include <stdio.h>

int main(int argc, char **argv)
{
    struct timespec elapsed_from_boot;

    clock_gettime(CLOCK_BOOTTIME, &elapsed_from_boot);

    printf("%d - seconds elapsed from boot\n", elapsed_from_boot.tv_sec);

    return 0;
}
```

which prints `uptime` information:

```
~$ gcc uptime.c -o uptime
~$ ./uptime
14180 - seconds elapsed from boot
```

We can easily check the result with the help of the uptime util:

```
~$ uptime
up  3:56
```

The `elapsed_from_boot.tv_sec` represents elapsed time in seconds, so:

```
>>> 14180 / 60
236
>>> 14180 / 60 / 60
3
>>> 14180 / 60 % 60
56
```

The `clock_id` maybe one of the following:

- `CLOCK_REALTIME` - system wide clock which measures real or wall-clock time;
- `CLOCK_REALTIME_COARSE` - faster version of the `CLOCK_REALTIME` ;
- `CLOCK_MONOTONIC` - represents monotonic time since some unspecified starting point;
- `CLOCK_MONOTONIC_COARSE` - faster version of the `CLOCK_MONOTONIC` ;
- `CLOCK_MONOTONIC_RAW` - the same as the `CLOCK_MONOTONIC` but provides non NTP adjusted time.
- `CLOCK_BOOTTIME` - the same as the `CLOCK_MONOTONIC` but plus time that the system was suspended;
- `CLOCK_PROCESS_CPUTIME_ID` - per-process time consumed by all threads in the process;
- `CLOCK_THREAD_CPUTIME_ID` - thread-specific clock.

The `clock_gettime` is not usual syscall too, but as the `gettimeofday` , this system call is placed in the `vDSO` area. Entry of this system call is located in the same source code file - arch/x86/entry/vdso/vclock_gettime.c) as for `gettimeofday` .

The Implementation of the `clock_gettime` depends on the clock id. If we have passed the `CLOCK_REALTIME` clock id, the `do_realtime` function will be called:

```
notrace int __vdso_clock_gettime(clockid_t clock, struct timespec *ts)
{
    switch (clock) {
    case CLOCK_REALTIME:
        if (do_realtime(ts) == VCLOCK_NONE)
            goto fallback;
        break;
    ...
    ...
    ...
fallback:
    return vdso_fallback_gettime(clock, ts);
}
```

In other cases, the `do_{name_of_clock_id}` function is called. Implementations of some of them is similar. For example if we will pass the `CLOCK_MONOTONIC` clock id:

```
...
...
...
case CLOCK_MONOTONIC:
    if (do_monotonic(ts) == VCLOCK_NONE)
        goto fallback;
    break;
...
...
...
```

the `do_monotonic` function will be called which is very similar on the implementation of the `do_realtime` :

```
notrace static int __always_inline do_monotonic(struct timespec *ts)
{
    do {
        seq = gtod_read_begin(gtod);
        mode = gtod->vclock_mode;
        ts->tv_sec = gtod->monotonic_time_sec;
        ns = gtod->monotonic_time_snsec;
        ns += vgetsns(&mode);
        ns >>= gtod->shift;
    } while (unlikely(gtod_read_retry(gtod, seq)));

    ts->tv_sec += __iter_div_u64_rem(ns, NSEC_PER_SEC, &ns);
    ts->tv_nsec = ns;

    return mode;
}
```

We already saw a little about the implementation of this function in the previous paragraph about the `gettimeofday`. There is only one difference here, that the `sec` and `nsec` of our `timespec` value will be based on the `gtod->monotonic_time_sec` instead of `gtod->wall_time_sec` which maps the value of the `tk->tkr_mono.xtime_nsec` or number of nanoseconds elapsed.

That's all.

Implementation of the `nanosleep` system call

The last system call in our list is the `nanosleep`. As you can understand from its name, this function provides `sleeping` ability. Let's look on the following simple example:

```
#include <time.h>
#include <stdlib.h>
#include <stdio.h>

int main (void)
{
   struct timespec ts = {5,0};

   printf("sleep five seconds\n");
   nanosleep(&ts, NULL);
   printf("end of sleep\n");

   return 0;
}
```

If we will compile and run it, we will see the first line

```
~$ gcc sleep_test.c -o sleep
~$ ./sleep
sleep five seconds
end of sleep
```

and the second line after five seconds.

The `nanosleep` is not located in the `vDSO` area like the `gettimeofday` and the `clock_gettime` functions. So, let's look how the `real` system call which is located in the kernel space will be called by the standard library. The implementation of the `nanosleep` system call will be called with the help of the syscall instruction. Before the execution of the `syscall` instruction, parameters of the system call must be put in processor registers according to order which is described in the System V Application Binary Interface or in other words:

- `rdi` - first parameter;

- `rsi` - second parameter;
- `rdx` - third parameter;
- `r10` - fourth parameter;
- `r8` - fifth parameter;
- `r9` - sixth parameter.

The `nanosleep` system call has two parameters - two pointers to the `timespec` structures. The system call suspends the calling thread until the given timeout has elapsed. Additionally it will finish if a signal interrupts its execution. It takes two parameters, the first is `timespec` which represents timeout for the sleep. The second parameter is the pointer to the `timespec` structure too and it contains remainder of time if the call of the `nanosleep` was interrupted.

As `nanosleep` has two parameters:

```
int nanosleep(const struct timespec *req, struct timespec *rem);
```

To call system call, we need put the `req` to the `rdi` register, and the `rem` parameter to the `rsi` register. The glibc does these job in the `INTERNAL_SYSCALL` macro which is located in the sysdeps/unix/sysv/linux/x86_64/sysdep.h header file.

```
# define INTERNAL_SYSCALL(name, err, nr, args...) \
  INTERNAL_SYSCALL_NCS (__NR_##name, err, nr, ##args)
```

which takes the name of the system call, storage for possible error during execution of system call, number of the system call (all `x86_64` system calls you can find in the system calls table) and arguments of certain system call. The `INTERNAL_SYSCALL` macro just expands to the call of the `INTERNAL_SYSCALL_NCS` macro, which prepares arguments of system call (puts them into the processor registers in correct order), executes `syscall` instruction and returns the result:

```
# define INTERNAL_SYSCALL_NCS(name, err, nr, args...)      \
  ({                                                        \
    unsigned long int resultvar;                            \
    LOAD_ARGS_##nr (args)                                    \
    LOAD_REGS_##nr                                            \
    asm volatile (                                            \
    "syscall\n\t"                                            \
    : "=a" (resultvar)                                        \
    : "0" (name) ASM_ARGS_##nr : "memory", REGISTERS_CLOBBERED_BY_SYSCALL);   \
    (long int) resultvar; })
```

The `LOAD_ARGS_##nr` macro calls the `LOAD_ARGS_N` macro where the `N` is number of arguments of the system call. In our case, it will be the `LOAD_ARGS_2` macro. Ultimately all of these macros will be expanded to the following:

```
# define LOAD_REGS_TYPES_1(t1, a1)                          \
  register t1 _a1 asm ("rdi") = __arg1;                      \
  LOAD_REGS_0

# define LOAD_REGS_TYPES_2(t1, a1, t2, a2)                  \
  register t2 _a2 asm ("rsi") = __arg2;                      \
  LOAD_REGS_TYPES_1(t1, a1)
...
...
...
```

After the `syscall` instruction will be executed, the context switch will occur and the kernel will transfer execution to the system call handler. The system call handler for the `nanosleep` system call is located in the kernel/time/hrtimer.c source code file and defined with the `SYSCALL_DEFINE2` macro helper:

```
SYSCALL_DEFINE2(nanosleep, struct timespec __user *, rqtp,
        struct timespec __user *, rmtp)
```

```
{
    struct timespec tu;

    if (copy_from_user(&tu, rqtp, sizeof(tu)))
        return -EFAULT;

    if (!timespec_valid(&tu))
        return -EINVAL;

    return hrtimer_nanosleep(&tu, rmtp, HRTIMER_MODE_REL, CLOCK_MONOTONIC);
}
```

More about the `SYSCALL_DEFINE2` macro you may read in the chapter about system calls. If we look at the implementation of the `nanosleep` system call, first of all we will see that it starts from the call of the `copy_from_user` function. This function copies the given data from the userspace to kernelspace. In our case we copy timeout value to sleep to the kernelspace `timespec` structure and check that the given `timespec` is valid by the call of the `timesc_valid` function:

```
static inline bool timespec_valid(const struct timespec *ts)
{
    if (ts->tv_sec < 0)
        return false;
    if ((unsigned long)ts->tv_nsec >= NSEC_PER_SEC)
        return false;
    return true;
}
```

which just checks that the given `timespec` does not represent date before `1970` and nanoseconds does not overflow `1` second. The `nanosleep` function ends with the call of the `hrtimer_nanosleep` function from the same source code file. The `hrtimer_nanosleep` function creates a timer and calls the `do_nanosleep` function. The `do_nanosleep` does main job for us. This function provides loop:

```
do {
    set_current_state(TASK_INTERRUPTIBLE);
    hrtimer_start_expires(&t->timer, mode);

    if (likely(t->task))
        freezable_schedule();

} while (t->task && !signal_pending(current));

__set_current_state(TASK_RUNNING);
return t->task == NULL;
```

Which freezes current task during sleep. After we set `TASK_INTERRUPTIBLE` flag for the current task, the `hrtimer_start_expires` function starts the give high-resolution timer on the current processor. As the given high resolution timer will expire, the task will be again running.

That's all.

Conclusion

This is the end of the seventh part of the chapter that describes timers and timer management related stuff in the Linux kernel. In the previous part we saw x86_64 specific clock sources. As I wrote in the beginning, this part is the last part of this chapter. We saw important time management related concepts like `clocksource` and `clockevents` frameworks, `jiffies` counter and etc., in this chpater. Of course this does not cover all of the time management in the Linux kernel. Many parts of this mostly related to the scheduling which we will see in other chapter.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- system call
- C programming language
- standard library
- glibc
- real time clock
- NTP
- nanoseconds
- register
- System V Application Binary Interface
- context switch
- Introduction to timers in the Linux kernel
- uptime
- system calls table for x86_64
- High Precision Event Timer
- Time Stamp Counter
- x86_64
- previous part

Synchronization primitives in the Linux kernel.

This chapter describes synchronization primitives in the Linux kernel.

- Introduction to spinlocks - the first part of this chapter describes implementation of spinlock mechanism in the Linux kernel.
- Queued spinlocks - the second part describes another type of spinlocks - queued spinlocks.
- Semaphores - this part describes implementation of `semaphore` synchronization primitive in the Linux kernel.
- Mutual exclusion - this part describes - `mutex` in the Linux kernel.
- Reader/Writer semaphores - this part describes special type of semaphores - `reader/writer` semaphores.
- Sequential locks - this part describes sequential locks in the Linux kernel.

Synchronization primitives in the Linux kernel. Part 1.

Introduction

This part opens a new chapter in the linux-insides book. Timers and time management related stuff was described in the previous chapter. Now time to go next. As you may understand from the part's title, this chapter will describe synchronization primitives in the Linux kernel.

As always, before we will consider something synchronization related, we will try to know what is `synchronization primitive` in general. Actually, synchronization primitive is a software mechanism which provides the ability to two or more parallel processes or threads to not execute simultaneously on the same segment of a code. For example, let's look on the following piece of code:

```
mutex_lock(&clocksource_mutex);
...
...
...
clocksource_enqueue(cs);
clocksource_enqueue_watchdog(cs);
clocksource_select();
...
...
...
mutex_unlock(&clocksource_mutex);
```

from the kernel/time/clocksource.c source code file. This code is from the `__clocksource_register_scale` function which adds the given clocksource to the clock sources list. This function produces different operations on a list with registered clock sources. For example, the `clocksource_enqueue` function adds the given clock source to the list with registered clocksources - `clocksource_list`. Note that these lines of code wrapped to two functions: `mutex_lock` and `mutex_unlock` which takes one parameter - the `clocksource_mutex` in our case.

These functions represent locking and unlocking based on mutex synchronization primitive. As `mutex_lock` will be executed, it allows us to prevent the situation when two or more threads will execute this code while the `mutex_unlock` will not be executed by process-owner of the mutex. In other words, we prevent parallel operations on a `clocksource_list`. Why do we need `mutex` here? What if two parallel processes will try to register a clock source. As we already know, the `clocksource_enqueue` function adds the given clock source to the `clocksource_list` list right after a clock source in the list which has the biggest rating (a registered clock source which has the highest frequency in the system):

```
static void clocksource_enqueue(struct clocksource *cs)
{
    struct list_head *entry = &clocksource_list;
    struct clocksource *tmp;

    list_for_each_entry(tmp, &clocksource_list, list)
        if (tmp->rating >= cs->rating)
            entry = &tmp->list;
    list_add(&cs->list, entry);
}
```

If two parallel processes will try to do it simultaneously, both process may found the same `entry` may occur race condition or in other words, the second process which will execute `list_add`, will overwrite a clock source from the first thread.

Besides this simple example, synchronization primitives are ubiquitous in the Linux kernel. If we will go through the previous chapter or other chapters again or if we will look at the Linux kernel source code in general, we will meet many places like this. We will not consider how `mutex` is implemented in the Linux kernel. Actually, the Linux kernel provides a set of different synchronization primitives like:

- `mutex` ;
- `semaphores` ;
- `seqlocks` ;
- `atomic operations` ;
- etc.

We will start this chapter from the `spinlock` .

Spinlocks in the Linux kernel.

The `spinlock` is a low-level synchronization mechanism which in simple words, represents a variable which can be in two states:

- `acquired` ;
- `released` .

Each process which wants to acquire a `spinlock` , must write a value which represents `spinlock acquired` state to this variable and write `spinlock released` state to the variable. If a process tries to execute code which is protected by a `spinlock` , it will be locked while a process which holds this lock will release it. In this case all related operations must be atomic to prevent race conditions state. The `spinlock` is represented by the `spinlock_t` type in the Linux kernel. If we will look at the Linux kernel code, we will see that this type is widely used. The `spinlock_t` is defined as:

```
typedef struct spinlock {
        union {
              struct raw_spinlock rlock;

#ifdef CONFIG_DEBUG_LOCK_ALLOC
# define LOCK_PADSIZE (offsetof(struct raw_spinlock, dep_map))
                struct {
                        u8 __padding[LOCK_PADSIZE];
                        struct lockdep_map dep_map;
                };
#endif
        };
} spinlock_t;
```

and located in the include/linux/spinlock_types.h header file. We may see that its implementation depends on the state of the `CONFIG_DEBUG_LOCK_ALLOC` kernel configuration option. We will skip this now, because all debugging related stuff will be in the end of this part. So, if the `CONFIG_DEBUG_LOCK_ALLOC` kernel configuration option is disabled, the `spinlock_t` contains union with one field which is - `raw_spinlock` :

```
typedef struct spinlock {
        union {
              struct raw_spinlock rlock;
        };
} spinlock_t;
```

The `raw_spinlock` structure defined in the same header file and represents the implementation of `normal` spinlock. Let's look how the `raw_spinlock` structure is defined:

```
typedef struct raw_spinlock {
```

```
        arch_spinlock_t raw_lock;
#ifdef CONFIG_GENERIC_LOCKBREAK
        unsigned int break_lock;
#endif
} raw_spinlock_t;
```

where the `arch_spinlock_t` represents architecture-specific `spinlock` implementation and the `break_lock` field which holds value - `1` in a case when one processor starts to wait while the lock is held on another processor on SMP systems. This allows prevent long time locking. As consider the x86_64 architecture in this books, so the `arch_spinlock_t` is defined in the arch/x86/include/asm/spinlock_types.h header file and looks:

```
#ifdef CONFIG_QUEUED_SPINLOCKS
#include <asm-generic/qspinlock_types.h>
#else
typedef struct arch_spinlock {
        union {
                __ticketpair_t head_tail;
                struct __raw_tickets {
                        __ticket_t head, tail;
                } tickets;
        };
} arch_spinlock_t;
```

As we may see, the definition of the `arch_spinlock` structure depends on the value of the `CONFIG_QUEUED_SPINLOCKS` kernel configuration option. This configuration option the Linux kernel supports `spinlocks` with queue. This special type of `spinlocks` which instead of `acquired` and `released` atomic values used `atomic` operation on a `queue` . If the `CONFIG_QUEUED_SPINLOCKS` kernel configuration option is enabled, the `arch_spinlock_t` will be represented by the following structure:

```
typedef struct qspinlock {
    atomic_t    val;
} arch_spinlock_t;
```

from the include/asm-generic/qspinlock_types.h header file.

We will not stop on this structures for now and before we will consider both `arch_spinlock` and the `qspinlock` , let's look at the operations on a spinlock. The Linux kernel provides following main operations on a `spinlock` :

- `spin_lock_init` - produces initialization of the given `spinlock` ;
- `spin_lock` - acquires given `spinlock` ;
- `spin_lock_bh` - disables software interrupts and acquire given `spinlock` .
- `spin_lock_irqsave` and `spin_lock_irq` - disable interrupts on local processor and preserve/not preserve previous interrupt state in the `flags` ;
- `spin_unlock` - releases given `spinlock` ;
- `spin_unlock_bh` - releases given `spinlock` and enables software interrupts;
- `spin_is_locked` - returns the state of the given `spinlock` ;
- and etc.

Let's look on the implementation of the `spin_lock_init` macro. As I already wrote, this and other macro are defined in the include/linux/spinlock.h header file and the `spin_lock_init` macro looks:

```
#define spin_lock_init(_lock)        \
do {                                            \
    spinlock_check(_lock);                        \
    raw_spin_lock_init(&(_lock)->rlock);        \
} while (0)
```

As we may see, the `spin_lock_init` macro takes a `spinlock` and executes two operations: check the given `spinlock` and execute the `raw_spin_lock_init`. The implementation of the `spinlock_check` is pretty easy, this function just returns the `raw_spinlock_t` of the given `spinlock` to be sure that we got exactly `normal` raw spinlock:

```
static __always_inline raw_spinlock_t *spinlock_check(spinlock_t *lock)
{
    return &lock->rlock;
}
```

The `raw_spin_lock_init` macro:

```
# define raw_spin_lock_init(lock)        \
do {                                             \
    *(lock) = __RAW_SPIN_LOCK_UNLOCKED(lock);         \
} while (0)                                      \
```

assigns the value of the `__RAW_SPIN_LOCK_UNLOCKED` with the given `spinlock` to the given `raw_spinlock_t`. As we may understand from the name of the `__RAW_SPIN_LOCK_UNLOCKED` macro, this macro does initialization of the given `spinlock` and set it to `released` state. This macro defined in the include/linux/spinlock_types.h header file and expands to the following macros:

```
#define __RAW_SPIN_LOCK_UNLOCKED(lockname)      \
         (raw_spinlock_t) __RAW_SPIN_LOCK_INITIALIZER(lockname)

#define __RAW_SPIN_LOCK_INITIALIZER(lockname)   \
         {                                               \
             .raw_lock = __ARCH_SPIN_LOCK_UNLOCKED,             \
             SPIN_DEBUG_INIT(lockname)                     \
             SPIN_DEP_MAP_INIT(lockname)                   \
         }
```

As I already wrote above, we will not consider stuff which is related to debugging of synchronization primitives. In this case we will not consider the `SPIN_DEBUG_INIT` and the `SPIN_DEP_MAP_INIT` macros. So the `__RAW_SPINLOCK_UNLOCKED` macro will be expanded to the:

```
*(&(_lock)->rlock) = __ARCH_SPIN_LOCK_UNLOCKED;
```

where the `__ARCH_SPIN_LOCK_UNLOCKED` is:

```
#define __ARCH_SPIN_LOCK_UNLOCKED       { { 0 } }
```

and:

```
#define __ARCH_SPIN_LOCK_UNLOCKED       { ATOMIC_INIT(0) }
```

for the x86_64 architecture. if the `CONFIG_QUEUED_SPINLOCKS` kernel configuration option is enabled. So, after the expansion of the `spin_lock_init` macro, a given `spinlock` will be initialized and its state will be - `unlocked`.

From this moment we know how to initialize a `spinlock`, now let's consider API which Linux kernel provides for manipulations of `spinlocks`. The first is:

```
static __always_inline void spin_lock(spinlock_t *lock)
{
    raw_spin_lock(&lock->rlock);
}
```

function which allows us to `acquire` a spinlock. The `raw_spin_lock` macro is defined in the same header file and expands to the call of the `_raw_spin_lock` function:

```
#define raw_spin_lock(lock)    _raw_spin_lock(lock)
```

As we may see in the include/linux/spinlock.h header file, definition of the `_raw_spin_lock` macro depends on the `CONFIG_SMP` kernel configuration parameter:

```
#if defined(CONFIG_SMP) || defined(CONFIG_DEBUG_SPINLOCK)
# include <linux/spinlock_api_smp.h>
#else
# include <linux/spinlock_api_up.h>
#endif
```

So, if the SMP is enabled in the Linux kernel, the `_raw_spin_lock` macro is defined in the arch/x86/include/asm/spinlock.h header file and looks like:

```
#define _raw_spin_lock(lock) __raw_spin_lock(lock)
```

The `__raw_spin_lock` function looks:

```
static inline void __raw_spin_lock(raw_spinlock_t *lock)
{
        preempt_disable();
        spin_acquire(&lock->dep_map, 0, 0, _RET_IP_);
        LOCK_CONTENDED(lock, do_raw_spin_trylock, do_raw_spin_lock);
}
```

As you may see, first of all we disable preemption by the call of the `preempt_disable` macro from the include/linux/preempt.h (more about this you may read in the ninth part of the Linux kernel initialization process chapter). When we will unlock the given `spinlock`, preemption will be enabled again:

```
static inline void __raw_spin_unlock(raw_spinlock_t *lock)
{
        ...
        ...
        ...
        preempt_enable();
}
```

We need to do this while a process is spinning on a lock, other processes must be prevented to preempt the process which acquired a lock. The `spin_acquire` macro which through a chain of other macros expands to the call of the:

```
#define spin_acquire(l, s, t, i)                lock_acquire_exclusive(l, s, t, NULL, i)
#define lock_acquire_exclusive(l, s, t, n, i)           lock_acquire(l, s, t, 0, 1, n, i)
```

`lock_acquire` function:

```
void lock_acquire(struct lockdep_map *lock, unsigned int subclass,
                  int trylock, int read, int check,
                  struct lockdep_map *nest_lock, unsigned long ip)
{
         unsigned long flags;

         if (unlikely(current->lockdep_recursion))
                return;
```

```
        raw_local_irq_save(flags);
        check_flags(flags);

        current->lockdep_recursion = 1;
        trace_lock_acquire(lock, subclass, trylock, read, check, nest_lock, ip);
        __lock_acquire(lock, subclass, trylock, read, check,
                       irqs_disabled_flags(flags), nest_lock, ip, 0, 0);
        current->lockdep_recursion = 0;
        raw_local_irq_restore(flags);
}
```

As I wrote above, we will not consider stuff here which is related to debugging or tracing. The main point of the `lock_acquire` function is to disable hardware interrupts by the call of the `raw_local_irq_save` macro, because the given spinlock might be acquired with enabled hardware interrupts. In this way the process will not be preempted. Note that in the end of the `lock_acquire` function we will enable hardware interrupts again with the help of the `raw_local_irq_restore` macro. As you already may guess, the main work will be in the `__lock_acquire` function which is defined in the kernel/locking/lockdep.c source code file.

The `__lock_acquire` function looks big. We will try to understand what does this function do, but not in this part. Actually this function mostly related to the Linux kernel lock validator and it is not topic of this part. If we will return to the definition of the `__raw_spin_lock` function, we will see that it contains the following definition in the end:

```
LOCK_CONTENDED(lock, do_raw_spin_trylock, do_raw_spin_lock);
```

The `LOCK_CONTENDED` macro is defined in the include/linux/lockdep.h header file and just calls the given function with the given `spinlock` :

```
#define LOCK_CONTENDED(_lock, try, lock) \
        lock(_lock)
```

In our case, the `lock` is `do_raw_spin_lock` function from the include/linux/spinlock.h header file and the `_lock` is the given `raw_spinlock_t` :

```
static inline void do_raw_spin_lock(raw_spinlock_t *lock) __acquires(lock)
{
       __acquire(lock);
        arch_spin_lock(&lock->raw_lock);
}
```

The `__acquire` here is just sparse related macro and we are not interested in it in this moment. Location of the definition of the `arch_spin_lock` function depends on two things: the first is the architecture of the system and the second do we use `queued spinlocks` or not. In our case we consider only `x86_64` architecture, so the definition of the `arch_spin_lock` is represented as the macro from the include/asm-generic/qspinlock.h header file:

```
#define arch_spin_lock(l)               queued_spin_lock(l)
```

if we are using `queued spinlocks` . Or in other case, the `arch_spin_lock` function is defined in the arch/x86/include/asm/spinlock.h header file. Now we will consider only `normal spinlock` and information related to `queued spinlocks` we will see later. Let's look again on the definition of the `arch_spinlock` structure, to understand the implementation of the `arch_spin_lock` function:

```
typedef struct arch_spinlock {
        union {
               __ticketpair_t head_tail;
               struct __raw_tickets {
                       __ticket_t head, tail;
```

```
                } tickets;
        };
} arch_spinlock_t;
```

This variant of `spinlock` is called - `ticket spinlock`. As we may see, it consists from two parts. When lock is acquired, it increments a `tail` by one every time when a process wants to hold a `spinlock`. If the `tail` is not equal to `head`, the process will be locked, until values of these variables will not be equal. Let's look on the implementation of the `arch_spin_lock` function:

```
static __always_inline void arch_spin_lock(arch_spinlock_t *lock)
{
        register struct __raw_tickets inc = { .tail = TICKET_LOCK_INC };

        inc = xadd(&lock->tickets, inc);

        if (likely(inc.head == inc.tail))
                goto out;

        for (;;) {
                 unsigned count = SPIN_THRESHOLD;

                 do {
                       inc.head = READ_ONCE(lock->tickets.head);
                       if (__tickets_equal(inc.head, inc.tail))
                                goto clear_slowpath;
                        cpu_relax();
                 } while (--count);
                 __ticket_lock_spinning(lock, inc.tail);
         }
clear_slowpath:
        __ticket_check_and_clear_slowpath(lock, inc.head);
out:
        barrier();
}
```

At the beginning of the `arch_spin_lock` function we can initialization of the `__raw_tickets` structure with `tail` - `1`:

```
#define __TICKET_LOCK_INC       1
```

In the next line we execute xadd operation on the `inc` and `lock->tickets`. After this operation the `inc` will store value of the `tickets` of the given `lock` and the `tickets.tail` will be increased on `inc` or `1`. The `tail` value was increased on `1` which means that one process started to try to hold a lock. In the next step we do the check that checks that `head` and `tail` have the same value. If these values are equal, this means that nobody holds lock and we go to the `out` label. In the end of the `arch_spin_lock` function we may see the `barrier` macro which represents `barrier instruction` which guarantees that compiler will not change the order of operations that access memory (more about memory barriers you can read in the kernel documentation).

If one process held a lock and a second process started to execute the `arch_spin_lock` function, the `head` will not be `equal` to `tail`, because the `tail` will be greater than `head` on `1`. In this way, the process will occur in the loop. There will be comparison between `head` and the `tail` values at each loop iteration. If these values are not equal, the `cpu_relax` will be called which is just NOP instruction:

```
#define cpu_relax()     asm volatile("rep; nop")
```

and the next iteration of the loop will be started. If these values will be equal, this means that the process which held this lock, released this lock and the next process may acquire the lock.

The `spin_unlock` operation goes through the all macros/function as `spin_lock`, of course with `unlock` prefix. In the end the `arch_spin_unlock` function will be called. If we will look at the implementation of the `arch_spin_unlock` function, we will see that it increases `head` of the `lock tickets` list:

```
__add(&lock->tickets.head, TICKET_LOCK_INC, UNLOCK_LOCK_PREFIX);
```

In a combination of the `spin_lock` and `spin_unlock`, we get kind of queue where `head` contains an index number which maps currently executed process which holds a lock and the `tail` which contains an index number which maps last process which tried to hold the lock:

```
     +-------+       +-------+
     |       |       |       |
head |   7   | - - - |   7   | tail
     |       |       |       |
     +-------+       +-------+
                         |
                     +-------+
                     |       |
                     |   8   |
                     |       |
                     +-------+
                         |
                     +-------+
                     |       |
                     |   9   |
                     |       |
                     +-------+
```

That's all for now. We didn't cover `spinlock` API in full in this part, but I think that the main idea behind this concept must be clear now.

Conclusion

This concludes the first part covering synchronization primitives in the Linux kernel. In this part, we met first synchronization primitive `spinlock` provided by the Linux kernel. In the next part we will continue to dive into this interesting theme and will see other `synchronization` related stuff.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- Concurrent computing
- Synchronization
- Clocksource framework
- Mutex
- Race condition
- Atomic operations
- SMP
- x86_64
- Interrupts
- Preemption

- Linux kernel lock validator
- Sparse
- xadd instruction
- NOP
- Memory barriers
- Previous chapter

Synchronization primitives in the Linux kernel. Part 2.

Queued Spinlocks

This is the second part of the chapter which describes synchronization primitives in the Linux kernel and in the first part of this chapter we met the first - spinlock. We will continue to learn this synchronization primitive in this part. If you have read the previous part, you may remember that besides normal spinlocks, the Linux kernel provides special type of `spinlocks` - `queued spinlocks`. In this part we will try to understand what does this concept represent.

We saw API of `spinlock` in the previous part:

- `spin_lock_init` - produces initialization of the given `spinlock` ;
- `spin_lock` - acquires given `spinlock` ;
- `spin_lock_bh` - disables software interrupts and acquire given `spinlock` .
- `spin_lock_irqsave` and `spin_lock_irq` - disable interrupts on local processor and preserve/not preserve previous interrupt state in the `flags` ;
- `spin_unlock` - releases given `spinlock` ;
- `spin_unlock_bh` - releases given `spinlock` and enables software interrupts;
- `spin_is_locked` - returns the state of the given `spinlock` ;
- and etc.

And we know that all of these macro which are defined in the include/linux/spinlock.h header file will be expanded to the call of the functions with `arch_spin_.*` prefix from the arch/x86/include/asm/spinlock.h for the x86_64 architecture. If we will look at this header fill with attention, we will that these functions (`arch_spin_is_locked` , `arch_spin_lock` , `arch_spin_unlock` and etc) defined only if the `CONFIG_QUEUED_SPINLOCKS` kernel configuration option is disabled:

```
#ifdef CONFIG_QUEUED_SPINLOCKS
#include <asm/qspinlock.h>
#else
static __always_inline void arch_spin_lock(arch_spinlock_t *lock)
{
    ...
    ...
    ...
}
...
...
...
#endif
```

This means that the arch/x86/include/asm/qspinlock.h header file provides own implementation of these functions. Actually they are macros and they are located in other header file. This header file is - include/asm-generic/qspinlock.h. If we will look into this header file, we will find definition of these macros:

```
#define arch_spin_is_locked(l)          queued_spin_is_locked(l)
#define arch_spin_is_contended(l)       queued_spin_is_contended(l)
#define arch_spin_value_unlocked(l)     queued_spin_value_unlocked(l)
#define arch_spin_lock(l)               queued_spin_lock(l)
#define arch_spin_trylock(l)            queued_spin_trylock(l)
#define arch_spin_unlock(l)             queued_spin_unlock(l)
#define arch_spin_lock_flags(l, f)      queued_spin_lock(l)
#define arch_spin_unlock_wait(l)        queued_spin_unlock_wait(l)
```

Before we will consider how queued spinlocks and their API are implemented, we take a look on theoretical part at first.

Introduction to queued spinlocks

Queued spinlocks is a locking mechanism in the Linux kernel which is replacement for the standard `spinlocks`. At least this is true for the x86_64 architecture. If we will look at the following kernel configuration file - kernel/Kconfig.locks, we will see following configuration entries:

```
config ARCH_USE_QUEUED_SPINLOCKS
    bool

config QUEUED_SPINLOCKS
    def_bool y if ARCH_USE_QUEUED_SPINLOCKS
    depends on SMP
```

This means that the `CONFIG_QUEUED_SPINLOCKS` kernel configuration option will be enabled by default if the `ARCH_USE_QUEUED_SPINLOCKS` is enabled. We may see that the `ARCH_USE_QUEUED_SPINLOCKS` is enabled by default in the `x86_64` specific kernel configuration file - arch/x86/Kconfig:

```
config X86
    ...
    ...
    ...
    select ARCH_USE_QUEUED_SPINLOCKS
    ...
    ...
    ...
```

Before we will start to consider what is it queued spinlock concept, let's look on other types of `spinlocks`. For the start let's consider how `normal` spinlocks is implemented. Usually, implementation of `normal` spinlock is based on the test and set instruction. Principle of work of this instruction is pretty simple. This instruction writes a value to the memory location and returns old value from this memory location. Both of these operations are in atomic context i.e. this instruction is non-interruptible. So if the first thread started to execute this instruction, second thread will wait until the first processor will not finish. Basic lock can be built on top of this mechanism. Schematically it may look like this:

```
int lock(lock)
{
    while (test_and_set(lock) == 1)
        ;
    return 0;
}

int unlock(lock)
{
    lock=0;

    return lock;
}
```

The first thread will execute the `test_and_set` which will set the `lock` to `1`. When the second thread will call the `lock` function, it will spin in the `while` loop, until the first thread will not call the `unlock` function and the `lock` will be equal to `0`. This implementation is not very good for performance, because it has at least two problems. The first problem is that this implementation may be unfair and the thread from one processor may have long waiting time, even if it called the `lock` before other threads which are waiting for free lock too. The second problem is that all threads which want to acquire a lock, must to execute many `atomic` operations like `test_and_set` on a variable which is in shared memory. This leads to the cache invalidation as the cache of the processor will store `lock=1`, but the value of the `lock` in memory may be `1` after a thread will release this lock.

In the previous part we saw the second type of spinlock implementation - `ticket spinlock`. This approach solves the first problem and may guarantee order of threads which want to acquire a lock, but still has a second problem.

The topic of this part is `queued spinlocks`. This approach may help to solve both of these problems. The `queued spinlocks` allows to each processor to use its own memory location to spin. The basic principle of a queue-based spinlock can best be understood by studying a classic queue-based spinlock implementation called the MCS lock. Before we will look at implementation of the `queued spinlocks` in the Linux kernel, we will try to understand what is it `MCS` lock.

The basic idea of the `MCS` lock is in that as I already wrote in the previous paragraph, a thread spins on a local variable and each processor in the system has its own copy of these variable. In other words this concept is built on top of the per-cpu variables concept in the Linux kernel.

When the first thread wants to acquire a lock, it registers itself in the `queue` or in other words it will be added to the special `queue` and will acquire lock, because it is free for now. When the second thread will want to acquire the same lock before the first thread will release it, this thread adds its own copy of the lock variable into this `queue`. In this case the first thread will contain a `next` field which will point to the second thread. From this moment, the second thread will wait until the first thread will release its lock and notify `next` thread about this event. The first thread will be deleted from the `queue` and the second thread will be owner of a lock.

Schematically we can represent it like:

Empty queue:

```
+---------+
|         |
|  Queue  |
|         |
+---------+
```

First thread tries to acquire a lock:

```
+---------+     +----------------------------+
|         |     |                            |
|  Queue  |---->| First thread acquired lock |
|         |     |                            |
+---------+     +----------------------------+
```

Second thread tries to acquire a lock:

```
+---------+     +-----------------------------------------+     +-------------------------+
|         |     |                                         |     |                         |
|  Queue  |---->|  Second thread waits for first thread   |<----| First thread holds lock |
|         |     |                                         |     |                         |
+---------+     +-----------------------------------------+     +-------------------------+
```

Or the pseudocode:

```
void lock(...)
{
    lock.next = NULL;
    ancestor = put_lock_to_queue_and_return_ancestor(queue, lock);

    // if we have ancestor, the lock already acquired and we
    // need to wait until it will be released
    if (ancestor)
    {
        lock.locked = 1;
        ancestor.next = lock;
```

```
        while (lock.is_locked == true)
            ;
    }

    // in other way we are owner of the lock and may exit
}

void unlock(...)
{
    // do we need to notify somebody or we are alonw in the
    // queue?
    if (lock.next != NULL) {
        // the while loop from the lock() function will be
        // finished
        lock.next.is_locked = false;
        // delete ourself from the queue and exit
        ...
        ...
        ...
        return;
    }

    // So, we have no next threads in the queue to notify about
    // lock releasing event. Let's just put `0` to the lock, will
    // delete ourself from the queue and exit.
}
```

The idea is simple, but the implementation of the `queued spinlocks` is more complex than this pseudocode. As I already wrote above, the `queued spinlock` mechanism is planned to be replacement for `ticket spinlocks` in the Linux kernel. But as you may remember, the usual `spinlock` fit into `32-bit` word. But the `MCS` based lock does not fit to this size. As you may know `spinlock_t` type is widely used in the Linux kernel. In this case would have to rewrite a significant part of the Linux kernel, but this is unacceptable. Beside this, some kernel structures which contains a spinlock for protection can't grow. But anyway, implementation of the `queued spinlocks` in the Linux kernel based on this concept with some modifications which allows to fit it into `32` bits.

That's all about theory of the `queued spinlocks`, now let's consider how this mechanism is implemented in the Linux kernel. Implementation of the `queued spinlocks` looks more complex and tangled than implementation of `ticket spinlocks`, but the study with attention will lead to success.

API of queued spinlocks

Now we know a little about `queued spinlocks` from the theoretical side, time to see the implementation of this mechanism in the Linux kernel. As we saw above, the include/asm-generic/qspinlock.h header files provides a set of macro which are represent API for a spinlock acquiring, releasing and etc:

```
#define arch_spin_is_locked(l)          queued_spin_is_locked(l)
#define arch_spin_is_contended(l)       queued_spin_is_contended(l)
#define arch_spin_value_unlocked(l)     queued_spin_value_unlocked(l)
#define arch_spin_lock(l)               queued_spin_lock(l)
#define arch_spin_trylock(l)            queued_spin_trylock(l)
#define arch_spin_unlock(l)             queued_spin_unlock(l)
#define arch_spin_lock_flags(l, f)      queued_spin_lock(l)
#define arch_spin_unlock_wait(l)        queued_spin_unlock_wait(l)
```

All of these macros expand to the call of functions from the same header file. Additionally, we saw the `qspinlock` structure from the include/asm-generic/qspinlock_types.h header file which represents a queued spinlock in the Linux kernel:

```
typedef struct qspinlock {
    atomic_t    val;
```

```
} arch_spinlock_t;
```

As we may see, the `qspinlock` structure contains only one field - `val`. This field represents the state of a given `spinlock`. This `4` bytes field consists from following four parts:

* `0-7` - locked byte;
* `8` - pending bit;
* `16-17` - two bit index which represents entry of the `per-cpu` array of the `MCS` lock (will see it soon);
* `18-31` - contains number of processor which indicates tail of the queue.

and the `9-15` bytes are not used.

As we already know, each processor in the system has own copy of the lock. The lock is represented by the following structure:

```
struct mcs_spinlock {
       struct mcs_spinlock *next;
       int locked;
       int count;
};
```

from the kernel/locking/mcs_spinlock.h header file. The first field represents a pointer to the next thread in the `queue`. The second field represents the state of the current thread in the `queue`, where `1` is `lock` already acquired and `0` in other way. And the last field of the `mcs_spinlock` structure represents nested locks. To understand what is it nested lock, imagine situation when a thread acquired lock, but was interrupted by the hardware interrupt and an interrupt handler tries to take a lock too. For this case, each processor has not just copy of the `mcs_spinlock` structure but array of these structures:

```
static DEFINE_PER_CPU_ALIGNED(struct mcs_spinlock, mcs_nodes[4]);
```

This array allows to make four attempts of a lock acquisition for the four events in following contexts:

* normal task context;
* hardware interrupt context;
* software interrupt context;
* non-maskable interrupt context.

Now let's return to the `qspinlock` structure and the `API` of the `queued spinlocks`. Before we will move to consider `API` of `queued spinlocks`, notice the `val` field of the `qspinlock` structure has type - `atomic_t` which represents atomic variable or one operation at a time variable. So, all operations with this field will be atomic. For example let's look at the reading value of the `val` API:

```
static __always_inline int queued_spin_is_locked(struct qspinlock *lock)
{
    return atomic_read(&lock->val);
}
```

Ok, now we know data structures which represents queued spinlock in the Linux kernel and now time is to look at the implementation of the `main` function from the `queued spinlocks` API.

```
#define arch_spin_lock(l)               queued_spin_lock(l)
```

Yes, this function is - `queued_spin_lock`. As we may understand from the function's name, it allows to acquire lock by the thread. This function is defined in the include/asm-generic/qspinlock_types.h header file and its implementation looks:

```
static __always_inline void queued_spin_lock(struct qspinlock *lock)
{
```

```
        u32 val;

        val = atomic_cmpxchg_acquire(&lock->val, 0, _Q_LOCKED_VAL);
        if (likely(val == 0))
                return;
        queued_spin_lock_slowpath(lock, val);
}
```

Looks pretty easy, except the `queued_spin_lock_slowpath` function. We may see that it takes only one parameter. In our case this parameter will represent `queued spinlock` which will be locked. Let's consider the situation that `queue` with locks is empty for now and the first thread wanted to acquire lock. As we may see the `queued_spin_lock` function starts from the call of the `atomic_cmpxchg_acquire` macro. As you may guess from the name of this macro, it executes atomic CMPXCHG instruction which compares value of the second parameter (zero in our case) with the value of the first parameter (current state of the given spinlock) and if they are identical, it stores value of the `_Q_LOCKED_VAL` in the memory location which is pointed by the `&lock->val` and return the initial value from this memory location.

The `atomic_cmpxchg_acquire` macro is defined in the include/linux/atomic.h header file and expands to the call of the `atomic_cmpxchg` function:

```
#define  atomic_cmpxchg_acquire         atomic_cmpxchg
```

which is architecture specific. We consider x86_64 architecture, so in our case this header file will be arch/x86/include/asm/atomic.h and the implementation of the `atomic_cmpxchg` function is just returns the result of the `cmpxchg` macro:

```
static __always_inline int atomic_cmpxchg(atomic_t *v, int old, int new)
{
        return cmpxchg(&v->counter, old, new);
}
```

This macro is defined in the arch/x86/include/asm/cmpxchg.h header file and looks:

```
#define cmpxchg(ptr, old, new) \
    __cmpxchg(ptr, old, new, sizeof(*(ptr)))

#define __cmpxchg(ptr, old, new, size) \
    __raw_cmpxchg((ptr), (old), (new), (size), LOCK_PREFIX)
```

As we may see, the `cmpxchg` macro expands to the `__cpmxchg` macro with the almost the same set of parameters. New additional parameter is the size of the atomic value. The `__cmpxchg` macro adds `LOCK_PREFIX` and expands to the `__raw_cmpxchg` macro where `LOCK_PREFIX` just LOCK instruction. After all, the `__raw_cmpxchg` does all job for us:

```
#define __raw_cmpxchg(ptr, old, new, size, lock) \
({
    ...
    ...
    ...
    volatile u32 *__ptr = (volatile u32 *)(ptr);            \
    asm volatile(lock "cmpxchgl %2,%1"                      \
                 : "=a" (__ret), "+m" (*__ptr)              \
                 : "r" (__new), "" (__old)                  \
                 : "memory");                               \
    ...
    ...
    ...
})
```

After the `atomic_cmpxchg_acquire` macro will be executed, it returns the previous value of the memory location. Now only one thread tried to acquire a lock, so the `val` will be zero and we will return from the `queued_spin_lock` function:

```
val = atomic_cmpxchg_acquire(&lock->val, 0, _Q_LOCKED_VAL);
if (likely(val == 0))
    return;
```

From this moment, our first thread will hold a lock. Notice that this behavior differs from the behavior which was described in the `MCS` algorithm. The thread acquired lock, but we didn't add it to the `queue`. As I already wrote the implementation of `queued spinlocks` concept is based on the `MCS` algorithm in the Linux kernel, but in the same time it has some difference like this for optimization purpose.

So the first thread have acquired lock and now let's consider that the second thread tried to acquire the same lock. The second thread will start from the same `queued_spin_lock` function, but the `lock->val` will contain `1` or `_Q_LOCKED_VAL`, because first thread already holds lock. So, in this case the `queued_spin_lock_slowpath` function will be called. The `queued_spin_lock_slowpath` function is defined in the kernel/locking/qspinlock.c source code file and starts from the following checks:

```
void queued_spin_lock_slowpath(struct qspinlock *lock, u32 val)
{
    if (pv_enabled())
        goto queue;

    if (virt_spin_lock(lock))
        return;

    ...
    ...
    ...
}
```

which check the state of the `pvqspinlock`. The `pvqspinlock` is `queued spinlock` in paravirtualized environment. As this chapter is related only to synchronization primitives in the Linux kernel, we skip these and other parts which are not directly related to the topic of this chapter. After these checks we compare our value which represents lock with the value of the `_Q_PENDING_VAL` macro and do nothing while this is true:

```
if (val == _Q_PENDING_VAL) {
    while ((val = atomic_read(&lock->val)) == _Q_PENDING_VAL)
        cpu_relax();
}
```

where `cpu_relax` is just NOP instruction. Above, we saw that the lock contains - `pending` bit. This bit represents thread which wanted to acquire lock, but it is already acquired by the other thread and in the same time `queue` is empty. In this case, the `pending` bit will be set and the `queue` will not be touched. This is done for optimization, because there are no need in unnecessary latency which will be caused by the cache invalidation in a touching of own `mcs_spinlock` array.

At the next step we enter into the following loop:

```
for (;;) {
    if (val & ~_Q_LOCKED_MASK)
        goto queue;

    new = _Q_LOCKED_VAL;
    if (val == new)
        new |= _Q_PENDING_VAL;

    old = atomic_cmpxchg_acquire(&lock->val, val, new);
    if (old == val)
```

```
        break;

    val = old;
}
```

The first `if` clause here checks that state of the lock (`val`) is in locked or pending state. This means that first thread already acquired lock, second thread tried to acquire lock too, but now it is in pending state. In this case we need to start to build queue. We will consider this situation little later. In our case we are first thread holds lock and the second thread tries to do it too. After this check we create new lock in a locked state and compare it with the state of the previous lock. As you remember, the `val` contains state of the `&lock->val` which after the second thread will call the `atomic_cmpxchg_acquire` macro will be equal to `1` . Both `new` and `val` values are equal so we set pending bit in the lock of the second thread. After this we need to check value of the `&lock->val` again, because the first thread may release lock before this moment. If the first thread did not released lock yet, the value of the `old` will be equal to the value of the `val` (because `atomic_cmpxchg_acquire` will return the value from the memory location which is pointed by the `lock->val` and now it is `1`) and we will exit from the loop. As we exited from this loop, we are waiting for the first thread until it will release lock, clear pending bit, acquire lock and return:

```
smp_cond_acquire(!(atomic_read(&lock->val) & _Q_LOCKED_MASK));
clear_pending_set_locked(lock);
return;
```

Notice that we did not touch `queue` yet. We no need in it, because for two threads it just leads to unnecessary latency for memory access. In other case, the first thread may release it lock before this moment. In this case the `lock->val` will contain `_Q_LOCKED_VAL | _Q_PENDING_VAL` and we will start to build `queue` . We start to build `queue` by the getting the local copy of the `mcs_nodes` array of the processor which executes thread:

```
node = this_cpu_ptr(&mcs_nodes[0]);
idx = node->count++;
tail = encode_tail(smp_processor_id(), idx);
```

Additionally we calculate `tail` which will indicate the tail of the `queue` and `index` which represents an entry of the `mcs_nodes` array. After this we set the `node` to point to the correct of the `mcs_nodes` array, set `locked` to zero because this thread didn't acquire lock yet and `next` to `NULL` because we don't know anything about other `queue` entries:

```
node += idx;
node->locked = 0;
node->next = NULL;
```

We already touch `per-cpu` copy of the queue for the processor which executes current thread which wants to acquire lock, this means that owner of the lock may released it before this moment. So we may try to acquire lock again by the call of the `queued_spin_trylock` function.

```
if (queued_spin_trylock(lock))
        goto release;
```

The `queued_spin_trylock` function is defined in the include/asm-generic/qspinlock.h header file and just does the same `queued_spin_lock` function that does:

```
static __always_inline int queued_spin_trylock(struct qspinlock *lock)
{
    if (!atomic_read(&lock->val) &&
       (atomic_cmpxchg_acquire(&lock->val, 0, _Q_LOCKED_VAL) == 0))
        return 1;
    return 0;
}
```

If the lock was successfully acquired we jump to the `release` label to release a node of the `queue` :

```
release:
    this_cpu_dec(mcs_nodes[0].count);
```

because we no need in it anymore as lock is acquired. If the `queued_spin_trylock` was unsuccessful, we update tail of the queue:

```
old = xchg_tail(lock, tail);
```

and retrieve previous tail. The next step is to check that `queue` is not empty. In this case we need to link previous entry with the new:

```
if (old & _Q_TAIL_MASK) {
    prev = decode_tail(old);
    WRITE_ONCE(prev->next, node);

    arch_mcs_spin_lock_contended(&node->locked);
}
```

After queue entries linked, we start to wait until reaching the head of queue. As we As we reached this, we need to do a check for new node which might be added during this wait:

```
next = READ_ONCE(node->next);
if (next)
    prefetchw(next);
```

If the new node was added, we prefetch cache line from memory pointed by the next queue entry with the PREFETCHW instruction. We preload this pointer now for optimization purpose. We just became a head of queue and this means that there is upcoming `MCS` unlock operation and the next entry will be touched.

Yes, from this moment we are in the head of the `queue` . But before we are able to acquire a lock, we need to wait at least two events: current owner of a lock will release it and the second thread with `pending` bit will acquire a lock too:

```
smp_cond_acquire(!((val = atomic_read(&lock->val)) & _Q_LOCKED_PENDING_MASK));
```

After both threads will release a lock, the head of the `queue` will hold a lock. In the end we just need to update the tail of the `queue` and remove current head from it.

That's all.

Conclusion

This is the end of the second part of the synchronization primitives chapter in the Linux kernel. In the previous part we already met the first synchronization primitive `spinlock` provided by the Linux kernel which is implemented as `ticket spinlock` . In this part we saw another implementation of the `spinlock` mechanism - `queued spinlock` . In the next part we will continue to dive into synchronization primitives in the Linux kernel.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

* spinlock
* interrupt
* interrupt handler
* API
* Test and Set
* MCS
* per-cpu variables
* atomic instruction
* CMPXCHG instruction
* LOCK instruction
* NOP instruction
* PREFETCHW instruction
* x86_64
* Previous part

Synchronization primitives in the Linux kernel. Part 3.

Semaphores

This is the third part of the chapter which describes synchronization primitives in the Linux kernel and in the previous part we saw special type of spinlocks - `queued spinlocks`. The previous part was the last part which describes `spinlocks` related stuff. So we need to go ahead.

The next synchronization primitive after `spinlock` which we will see in this part is semaphore. We will start from theoretical side and will learn what is it `semaphore` and only after this, we will see how it is implemented in the Linux kernel as we did in the previous part.

So, let's start.

Introduction to the semaphores in the Linux kernel

So, what is it `semaphore`? As you may guess - `semaphore` is yet another mechanism for support of thread or process synchronization. The Linux kernel already provides implementation of one synchronization mechanism - `spinlocks`, why do we need in yet another one? To answer on this question we need to know details of both of these mechanisms. We already familiar with the `spinlocks`, so let's start from this mechanism.

The main idea behind `spinlock` concept is a lock which will be acquired for a very short time. We can't sleep when a lock acquired by a process or thread, because other processes wait us. Context switch is not not allowed because preemption is disabled to avoid deadlocks.

In this way, semaphores is a good solution for locks which may be acquired for a long time. In other way this mechanism is not optimal for locks that acquired for a short time. To understand this, we need to know what is `semaphore`.

As usual synchronization primitive, a `semaphore` is based on a variable. This variable may be incremented or decremented and it's state will represent ability to acquire lock. Notice that value of the variable is not limited to `0` and `1`. There are two types of `semaphores`:

* `binary semaphore`;
* `normal semaphore`.

In the first case, value of `semaphore` may be only `1` or `0`. In the second case value of `semaphore` any non-negative number. If the value of `semaphore` is greater than `1` it is called as `counting semaphore` and it allows to acquire a lock to more than `1` process. This allows us to keep records of available resources, when `spinlock` allows to hold a lock only on one task. Besides all of this, one more important thing that `semaphore` allows to sleep. Moreover when processes waits for a lock which is acquired by other process, the scheduler may switch on another process.

Semaphore API

So, we know a little about `semaphores` from theoretical side, let's look on its implementation in the Linux kernel. All `semaphore` API is located in the include/linux/semaphore.h header file.

We may see that the `semaphore` mechanism is represented by the following structure:

```
struct semaphore {
	raw_spinlock_t		lock;
	unsigned int		count;
	struct list_head	wait_list;
```

```
};
```

in the Linux kernel. The `semaphore` structure consists of three fields:

* `lock` - `spinlock` for a `semaphore` data protection;
* `count` - amount available resources;
* `wait_list` - list of processes which are waiting to acquire a lock.

Before we will consider an API of the `semaphore` mechanism in the Linux kernel, we need to know how to initialize a `semaphore`. Actually the Linux kernel provides two approaches to execute initialization of the given `semaphore` structure. These methods allows to initialize a `semaphore` in a:

* `statically`;
* `dynamically`.

ways. Let's look at the first approach. We are able to initialize a `semaphore` statically with the `DEFINE_SEMAPHORE` macro:

```
#define DEFINE_SEMAPHORE(name)  \
         struct semaphore name = __SEMAPHORE_INITIALIZER(name, 1)
```

as we may see, the `DEFINE_SEMAPHORE` macro provides ability to initialize only `binary` semaphore. The `DEFINE_SEMAPHORE` macro expands to the definition of the `semaphore` structure which is initialized with the `__SEMAPHORE_INITIALIZER` macro. Let's look at the implementation of this macro:

```
#define __SEMAPHORE_INITIALIZER(name, n)              \
{                                                                  \
        .lock           = __RAW_SPIN_LOCK_UNLOCKED((name).lock),        \
        .count          = n,                                        \
        .wait_list      = LIST_HEAD_INIT((name).wait_list),             \
}
```

The `__SEMAPHORE_INITIALIZER` macro takes the name of the future `semaphore` structure and does initialization of the fields of this structure. First of all we initialize a `spinlock` of the given `semaphore` with the `__RAW_SPIN_LOCK_UNLOCKED` macro. As you may remember from the previous parts, the `__RAW_SPIN_LOCK_UNLOCKED` is defined in the include/linux/spinlock_types.h header file and expands to the `__ARCH_SPIN_LOCK_UNLOCKED` macro which just expands to zero or unlocked state:

```
#define __ARCH_SPIN_LOCK_UNLOCKED       { { 0 } }
```

The last two fields of the `semaphore` structure `count` and `wait_list` are initialized with the given value which represents count of available resources and empty list.

The second way to initialize a `semaphore` structure is to pass the `semaphore` and number of available resources to the `sema_init` function which is defined in the include/linux/semaphore.h header file:

```
static inline void sema_init(struct semaphore *sem, int val)
{
       static struct lock_class_key __key;
       *sem = (struct semaphore) __SEMAPHORE_INITIALIZER(*sem, val);
       lockdep_init_map(&sem->lock.dep_map, "semaphore->lock", &__key, 0);
}
```

Let's consider implementation of this function. It looks pretty easy and actually it does almost the same. This function executes initialization of the given `semaphore` with the `__SEMAPHORE_INITIALIZER` macro which we just saw. As I already wrote in the previous parts of this chapter, we will skip the stuff which is related to the lock validator of the Linux kernel.

So, from now we are able to initialize a `semaphore` let's look at how to lock and unlock. The Linux kernel provides following API to manipulate `semaphores` :

```
void down(struct semaphore *sem);
void up(struct semaphore *sem);
int  down_interruptible(struct semaphore *sem);
int  down_killable(struct semaphore *sem);
int  down_trylock(struct semaphore *sem);
int  down_timeout(struct semaphore *sem, long jiffies);
```

The first two functions: `down` and `up` are for acquiring and releasing of the given `semaphore` . The `down_interruptible` function tries to acquire a `semaphore` . If this try was successful, the `count` field of the given `semaphore` will be decremented and lock will be acquired, in other way the task will be switched to the blocked state or in other words the `TASK_INTERRUPTIBLE` flag will be set. This `TASK_INTERRUPTIBLE` flag means that the process may return to running state by a signal.

The `down_killable` function does the same as the `down_interruptible` function, but set the `TASK_KILLABLE` flag for the current process. This means that the waiting process may be interrupted by the kill signal.

The `down_trylock` function is similar on the `spin_trylock` function. This function tries to acquire a lock and exit if this operation was unsuccessful. In this case the process which wants to acquire a lock, will not wait. The last `down_timeout` function tries to acquire a lock. It will be interrupted in a waiting state when the given timeout will be expired. Additionally, you may notice that the timeout is in jiffies

We just saw definitions of the `semaphore` API. We will start from the `down` function. This function is defined in the kernel/locking/semaphore.c source code file. Let's look on the implementation function:

```
void down(struct semaphore *sem)
{
        unsigned long flags;

        raw_spin_lock_irqsave(&sem->lock, flags);
        if (likely(sem->count > 0))
                sem->count--;
        else
                __down(sem);
        raw_spin_unlock_irqrestore(&sem->lock, flags);
}
EXPORT_SYMBOL(down);
```

We may see the definition of the `flags` variable at the beginning of the `down` function. This variable will be passed to the `raw_spin_lock_irqsave` and `raw_spin_lock_irqrestore` macros which are defined in the include/linux/spinlock.h header file and protect a counter of the given `semaphore` here. Actually both of these macro do the same that `spin_lock` and `spin_unlock` macros, but additionally they save/restore current value of interrupt flags and disables interrupts.

As you already may guess, the main work is done between the `raw_spin_lock_irqsave` and `raw_spin_unlock_irqrestore` macros in the `down` function. We compare the value of the `semaphore` counter with zero and if it is bigger than zero, we may decrement this counter. This means that we already acquired the lock. In other way counter is zero. This means that all available resources already finished and we need to wait to acquire this lock. As we may see, the `__down` function will be called in this case.

The `__down` function is defined in the same source code file and its implementation looks:

```
static noinline void __sched __down(struct semaphore *sem)
{
        __down_common(sem, TASK_UNINTERRUPTIBLE, MAX_SCHEDULE_TIMEOUT);
}
```

The `__down` function just calls the `__down_common` function with three parameters:

- `semaphore` ;
- `flag` - for the task;
- `timeout` - maximum timeout to wait `semaphore` .

Before we will consider implementation of the `__down_common` function, notice that implementation of the `down_trylock` , `down_timeout` and `down_killable` functions based on the `__down_common` too:

```
static noinline int __sched __down_interruptible(struct semaphore *sem)
{
        return __down_common(sem, TASK_INTERRUPTIBLE, MAX_SCHEDULE_TIMEOUT);
}
```

The `__down_killable` :

```
static noinline int __sched __down_killable(struct semaphore *sem)
{
        return __down_common(sem, TASK_KILLABLE, MAX_SCHEDULE_TIMEOUT);
}
```

And the `__down_timeout` :

```
static noinline int __sched __down_timeout(struct semaphore *sem, long timeout)
{
        return __down_common(sem, TASK_UNINTERRUPTIBLE, timeout);
}
```

Now let's look at the implementation of the `__down_common` function. This function is defined in the kernel/locking/semaphore.c source code file too and starts from the definition of the two following local variables:

```
struct task_struct *task = current;
struct semaphore_waiter waiter;
```

The first represents current task for the local processor which wants to acquire a lock. The `current` is a macro which is defined in the arch/x86/include/asm/current.h header file:

```
#define current get_current()
```

Where the `get_current` function returns value of the `current_task` per-cpu variable:

```
DECLARE_PER_CPU(struct task_struct *, current_task);

static __always_inline struct task_struct *get_current(void)
{
        return this_cpu_read_stable(current_task);
}
```

The second variable is `waiter` represents an entry of a `semaphore.wait_list` list:

```
struct semaphore_waiter {
        struct list_head list;
        struct task_struct *task;
        bool up;
};
```

Next we add current task to the `wait_list` and fill `waiter` fields after definition of these variables:

```
list_add_tail(&waiter.list, &sem->wait_list);
waiter.task = task;
waiter.up = false;
```

In the next step we join into the following infinite loop:

```
for (;;) {
        if (signal_pending_state(state, task))
            goto interrupted;

        if (unlikely(timeout <= 0))
            goto timed_out;

        __set_task_state(task, state);

        raw_spin_unlock_irq(&sem->lock);
        timeout = schedule_timeout(timeout);
        raw_spin_lock_irq(&sem->lock);

        if (waiter.up)
            return 0;
}
```

In the previous piece of code we set `waiter.up` to `false`. So, a task will spin in this loop while `up` will not be set to `true`. This loop starts from the check that the current task is in the `pending` state or in other words flags of this task contains `TASK_INTERRUPTIBLE` or `TASK_WAKEKILL` flag. As I already wrote above a task may be interrupted by signal during wait of ability to acquire a lock. The `signal_pending_state` function is defined in the include/linux/sched.h source code file and looks:

```
static inline int signal_pending_state(long state, struct task_struct *p)
{
         if (!(state & (TASK_INTERRUPTIBLE | TASK_WAKEKILL)))
                 return 0;
         if (!signal_pending(p))
                 return 0;

         return (state & TASK_INTERRUPTIBLE) || __fatal_signal_pending(p);
}
```

We check that the `state` bitmask contains `TASK_INTERRUPTIBLE` or `TASK_WAKEKILL` bits and if the bitmask does not contain this bit we exit. At the next step we check that the given task has a pending signal and exit if there is no. In the end we just check `TASK_INTERRUPTIBLE` bit in the `state` bitmask again or the SIGKILL signal. So, if our task has a pending signal, we will jump at the `interrupted` label:

```
interrupted:
    list_del(&waiter.list);
    return -EINTR;
```

where we delete task from the list of lock waiters and return the `-EINTR` error code. If a task has no pending signal, we check the given timeout and if it is less or equal zero:

```
if (unlikely(timeout <= 0))
    goto timed_out;
```

we jump at the `timed_out` label:

```
timed_out:
    list_del(&waiter.list);
    return -ETIME;
```

Where we do almost the same that we did in the `interrupted` label. We delete task from the list of lock waiters, but return the `-ETIME` error code. If a task has no pending signal and the given timeout is not expired yet, the given `state` will be set in the given task:

```
__set_task_state(task, state);
```

and call the `schedule_timeout` function:

```
raw_spin_unlock_irq(&sem->lock);
timeout = schedule_timeout(timeout);
raw_spin_lock_irq(&sem->lock);
```

which is defined in the kernel/time/timer.c source code file. The `schedule_timeout` function makes the current task sleep until the given timeout.

That is all about the `__down_common` function. A task which wants to acquire a lock which is already acquired by another task will be spun in the infinite loop while it will not be interrupted by a signal, the given timeout will not be expired or the task which holds a lock will not release it. Now let's look at the implementation of the `up` function.

The `up` function is defined in the same source code file as `down` function. As we already know, the main purpose of this function is to release a lock. This function looks:

```
void up(struct semaphore *sem)
{
        unsigned long flags;

        raw_spin_lock_irqsave(&sem->lock, flags);
        if (likely(list_empty(&sem->wait_list)))
                sem->count++;
        else
                __up(sem);
        raw_spin_unlock_irqrestore(&sem->lock, flags);
}
EXPORT_SYMBOL(up);
```

It looks almost the same as the `down` function. There are only two differences here. First of all we increment a counter of a `semaphore` if the list of waiters is empty. In other way we call the `__up` function from the same source code file. If the list of waiters is not empty we need to allow the first task from the list to acquire a lock:

```
static noinline void __sched __up(struct semaphore *sem)
{
        struct semaphore_waiter *waiter = list_first_entry(&sem->wait_list,
                                                struct semaphore_waiter, list);
        list_del(&waiter->list);
        waiter->up = true;
        wake_up_process(waiter->task);
}
```

Here we takes the first task from the list of waiters, delete it from the list, set its `waiter-up` to true. From this point the infinite loop from the `__down_common` function will be stopped. The `wake_up_process` function will be called in the end of the `__up` function. As you remember we called the `schedule_timeout` function in the infinite loop from the `__down_common` this function. The `schedule_timeout` function makes the current task sleep until the given timeout will not be expired. So, as our process may sleep right now, we need to wake it up. That's why we call the `wake_up_process` function from the kernel/sched/core.c source code file.

That's all.

Conclusion

This is the end of the third part of the synchronization primitives chapter in the Linux kernel. In the two previous parts we already met the first synchronization primitive `spinlock` provided by the Linux kernel which is implemented as `ticket spinlock` and used for a very short time locks. In this part we saw yet another synchronization primitive - semaphore which is used for long time locks as it leads to context switch. In the next part we will continue to dive into synchronization primitives in the Linux kernel and will see next synchronization primitive - mutex.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- spinlocks
- synchronization primitive
- semaphore
- context switch
- preemption
- deadlocks
- scheduler
- Doubly linked list in the Linux kernel
- jiffies
- interrupts
- per-cpu
- bitmask
- SIGKILL
- errno
- API
- mutex
- Previous part

Synchronization primitives in the Linux kernel. Part 4.

Introduction

This is the fourth part of the chapter which describes synchronization primitives in the Linux kernel and in the previous parts we finished to consider different types spinlocks and semaphore synchronization primitives. We will continue to learn synchronization primitives in this part and consider yet another one which is called - mutex which is stands for `MUTual EXclusion` .

As in all previous parts of this book, we will try to consider this synchronization primitive from the theoretical side and only than we will consider API provided by the Linux kernel to manipulate with `mutexes` .

So, let's start.

Concept of `mutex`

We already familiar with the semaphore synchronization primitive from the previous part. It represented by the:

```
struct semaphore {
    raw_spinlock_t        lock;
    unsigned int        count;
    struct list_head    wait_list;
};
```

structure which holds information about state of a lock and list of a lock waiters. Depends on the value of the `count` field, a `semaphore` can provide access to a resource of more than one wishing of this resource. The mutex concept is very similar to a semaphore concept. But it has some differences. The main difference between `semaphore` and `mutex` synchronization primitive is that `mutex` has more strict semantic. Unlike a `semaphore` , only one process may hold `mutex` at one time and only the `owner` of a `mutex` may release or unlock it. Additional difference in implementation of `lock` API. The `semaphore` synchronization primitive forces rescheduling of processes which are in waiters list. The implementation of `mutex` lock `API` allows to avoid this situation and as a result expensive context switches.

The `mutex` synchronization primitive represented by the following:

```
struct mutex {
        atomic_t                count;
        spinlock_t              wait_lock;
        struct list_head        wait_list;
#if defined(CONFIG_DEBUG_MUTEXES) || defined(CONFIG_MUTEX_SPIN_ON_OWNER)
        struct task_struct      *owner;
#endif
#ifdef CONFIG_MUTEX_SPIN_ON_OWNER
        struct optimistic_spin_queue osq;
#endif
#ifdef CONFIG_DEBUG_MUTEXES
        void                    *magic;
#endif
#ifdef CONFIG_DEBUG_LOCK_ALLOC
        struct lockdep_map      dep_map;
#endif
};
```

structure in the Linux kernel. This structure is defined in the include/linux/mutex.h header file and contains similar to the `semaphore` structure set of fields. The first field of the `mutex` structure is - `count` . Value of this field represents state of a `mutex` . In a case when the value of the `count` field is `1` , a `mutex` is in `unlocked` state. When the value of the `count` field is `zero` , a `mutex` is in the `locked` state. Additionally value of the `count` field may be `negative` . In this case a `mutex` is in the `locked` state and has possible waiters.

The next two fields of the `mutex` structure - `wait_lock` and `wait_list` are spinlock for the protection of a `wait queue` and list of waiters which represents this `wait queue` for a certain lock. As you may notice, the similarity of the `mutex` and `semaphore` structures ends. Remaining fields of the `mutex` structure, as we may see depends on different configuration options of the Linux kernel.

The first field - `owner` represents process which acquired a lock. As we may see, existence of this field in the `mutex` structure depends on the `CONFIG_DEBUG_MUTEXES` or `CONFIG_MUTEX_SPIN_ON_OWNER` kernel configuration options. Main point of this field and the next `osq` fields is support of `optimistic spinning` which we will see later. The last two fields - `magic` and `dep_map` are used only in debugging mode. The `magic` field is to storing a `mutex` related information for debugging and the second field - `lockdep_map` is for lock validator of the Linux kernel.

Now, after we have considered the `mutex` structure, we may consider how this synchronization primitive works in the Linux kernel. As you may guess, a process which wants to acquire a lock, must decrease value of the `mutex->count` if possible. And if a process wants to release a lock, it must increase the same value. That's true. But as you may also guess, it is not so simple in the Linux kernel.

Actually, when a process try to acquire a `mutex` , there three possible paths:

- `fastpath` ;
- `midpath` ;
- `slowpath` .

which may be taken, depending on the current state of the `mutex` . The first path or `fastpath` is the fastest as you may understand from its name. Everything is easy in this case. Nobody acquired a `mutex` , so the value of the `count` field of the `mutex` structure may be directly decremented. In a case of unlocking of a `mutex` , the algorithm is the same. A process just increments the value of the `count` field of the `mutex` structure. Of course, all of these operations must be atomic.

Yes, this looks pretty easy. But what happens if a process wants to acquire a `mutex` which is already acquired by other process? In this case, the control will be transferred to the second path - `midpath` . The `midpath` or `optimistic spinning` tries to spin with already familiar for us MCS lock while the lock owner is running. This path will be executed only if there are no other processes ready to run that have higher priority. This path is called `optimistic` because the waiting task will not sleep and will not be rescheduled. This allows to avoid expensive context switch.

In the last case, when the `fastpath` and `midpath` may not be executed, the last path - `slowpath` will be executed. This path acts like a semaphore lock. If the lock is unable to be acquired by a process, this process will be added to `wait queue` which is represented by the following:

```
struct mutex_waiter {
        struct list_head        list;
        struct task_struct      *task;
#ifdef CONFIG_DEBUG_MUTEXES
        void                    *magic;
#endif
};
```

structure from the include/linux/mutex.h header file and will be sleep. Before we will consider API which is provided by the Linux kernel for manipulation with `mutexes` , let's consider the `mutex_waiter` structure. If you have read the previous part of this chapter, you may notice that the `mutex_waiter` structure is similar to the `semaphore_waiter` structure from the kernel/locking/semaphore.c source code file:

```
struct semaphore_waiter {
        struct list_head list;
        struct task_struct *task;
        bool up;
};
```

It also contains `list` and `task` fields which are represent entry of the mutex wait queue. The one difference here that the `mutex_waiter` does not contain `up` field, but contains the `magic` field which depends on the `CONFIG_DEBUG_MUTEXES` kernel configuration option and used to store a `mutex` related information for debugging purpose.

Now we know what is it `mutex` and how it is represented the Linux kernel. In this case, we may go ahead and start to look at the API which the Linux kernel provides for manipulation of `mutexes` .

Mutex API

Ok, in the previous paragraph we knew what is it `mutex` synchronization primitive and saw the `mutex` structure which represents `mutex` in the Linux kernel. Now it's time to consider API for manipulation of mutexes. Description of the `mutex` API is located in the include/linux/mutex.h header file. As always, before we will consider how to acquire and release a `mutex` , we need to know how to initialize it.

There are two approaches to initialize a `mutex` . The first is to do it statically. For this purpose the Linux kernel provides following:

```
#define DEFINE_MUTEX(mutexname) \
        struct mutex mutexname = __MUTEX_INITIALIZER(mutexname)
```

macro. Let's consider implementation of this macro. As we may see, the `DEFINE_MUTEX` macro takes name for the `mutex` and expands to the definition of the new `mutex` structure. Additionally new `mutex` structure get initialized with the `__MUTEX_INITIALIZER` macro. Let's look at the implementation of the `__MUTEX_INITIALIZER` :

```
#define __MUTEX_INITIALIZER(lockname)         \
{                                                             \
       .count = ATOMIC_INIT(1),                               \
       .wait_lock = __SPIN_LOCK_UNLOCKED(lockname.wait_lock), \
       .wait_list = LIST_HEAD_INIT(lockname.wait_list)        \
}
```

This macro is defined in the same header file and as we may understand it initializes fields of the `mutex` structure the initial values. The `count` field get initialized with the `1` which represents `unlocked` state of a mutex. The `wait_lock` spinlock get initialized to the unlocked state and the last field `wait_list` to empty doubly linked list.

The second approach allows us to initialize a `mutex` dynamically. To do this we need to call the `__mutex_init` function from the kernel/locking/mutex.c source code file. Actually, the `__mutex_init` function rarely called directly. Instead of the `__mutex_init` , the:

```
# define mutex_init(mutex)                \
do {                                                  \
        static struct lock_class_key __key;           \
                                                      \
        __mutex_init((mutex), #mutex, &__key);        \
} while (0)
```

macro is used. We may see that the `mutex_init` macro just defines the `lock_class_key` and call the `__mutex_init` function. Let's look at the implementation of this function:

```
void
__mutex_init(struct mutex *lock, const char *name, struct lock_class_key *key)
{
        atomic_set(&lock->count, 1);
        spin_lock_init(&lock->wait_lock);
        INIT_LIST_HEAD(&lock->wait_list);
        mutex_clear_owner(lock);
#ifdef CONFIG_MUTEX_SPIN_ON_OWNER
        osq_lock_init(&lock->osq);
#endif
        debug_mutex_init(lock, name, key);
}
```

As we may see the `__mutex_init` function takes three arguments:

- `lock` - a mutex itself;
- `name` - name of mutex for debugging purpose;
- `key` - key for lock validator.

At the beginning of the `__mutex_init` function, we may see initialization of the `mutex` state. We set it to `unlocked` state with the `atomic_set` function which atomically set the give variable to the given value. After this we may see initialization of the `spinlock` to the unlocked state which will protect `wait queue` of the `mutex` and initialization of the `wait queue` of the `mutex`. After this we clear owner of the `lock` and initialize optimistic queue by the call of the `osq_lock_init` function from the include/linux/osq_lock.h header file. This function just sets the tail of the optimistic queue to the unlocked state:

```
static inline bool osq_is_locked(struct optimistic_spin_queue *lock)
{
        return atomic_read(&lock->tail) != OSQ_UNLOCKED_VAL;
}
```

In the end of the `__mutex_init` function we may see the call of the `debug_mutex_init` function, but as I already wrote in previous parts of this chapter, we will not consider debugging related stuff in this chapter.

After the `mutex` structure is initialized, we may go ahead and will look at the `lock` and `unlock` API of `mutex` synchronization primitive. Implementation of `mutex_lock` and `mutex_unlock` functions located in the kernel/locking/mutex.c source code file. First of all let's start from the implementation of the `mutex_lock`. It looks:

```
void __sched mutex_lock(struct mutex *lock)
{
        might_sleep();
        __mutex_fastpath_lock(&lock->count, __mutex_lock_slowpath);
        mutex_set_owner(lock);
}
```

We may see the call of the `might_sleep` macro from the include/linux/kernel.h header file at the beginning of the `mutex_lock` function. Implementation of this macro depends on the `CONFIG_DEBUG_ATOMIC_SLEEP` kernel configuration option and if this option is enabled, this macro just prints a stack trace if it was executed in atomic context. This macro is helper for debugging purposes. In other way this macro does nothing.

After the `might_sleep` macro, we may see the call of the `__mutex_fastpath_lock` function. This function is architecture-specific and as we consider x86_64 architecture in this book, the implementation of the `__mutex_fastpath_lock` is located in the arch/x86/include/asm/mutex_64.h header file. As we may understand from the name of the `__mutex_fastpath_lock` function, this function will try to acquire lock in a fast path or in other words this function will try to decrement the value of the `count` of the given mutex.

Implementation of the `__mutex_fastpath_lock` function consists from two parts. The first part is inline assembly statement. Let's look at it:

```
asm_volatile_goto(LOCK_PREFIX "   decl %0\n"
                              "   jns %l[exit]\n"
                              : : "m" (v->counter)
                              : "memory", "cc"
                              : exit);
```

First of all, let's pay attention to the `asm_volatile_goto`. This macro is defined in the include/linux/compiler-gcc.h header file and just expands to the two inline assembly statements:

```
#define asm_volatile_goto(x...) do { asm goto(x); asm (""); } while (0)
```

The first assembly statement contains `goto` specificator and the second empty inline assembly statement is barrier. Now let's return to the our inline assembly statement. As we may see it starts from the definition of the `LOCK_PREFIX` macro which just expands to the lock instruction:

```
#define LOCK_PREFIX LOCK_PREFIX_HERE "\n\tlock; "
```

As we already know from the previous parts, this instruction allows to execute prefixed instruction atomically. So, at the first step in the our assembly statement we try decrement value of the given `mutex->counter`. At the next step the jns instruction will execute jump at the `exit` label if the value of the decremented `mutex->counter` is not negative. The `exit` label is the second part of the `__mutex_fastpath_lock` function and it just points to the exit from this function:

```
exit:
        return;
```

For this moment he implementation of the `__mutex_fastpath_lock` function looks pretty easy. But the value of the `mutex->counter` may be negative after increment. In this case the:

```
fail_fn(v);
```

will be called after our inline assembly statement. The `fail_fn` is the second parameter of the `__mutex_fastpath_lock` function and represents pointer to function which represents `midpath/slowpath` paths to acquire the given lock. In our case the `fail_fn` is the `__mutex_lock_slowpath` function. Before we will look at the implementation of the `__mutex_lock_slowpath` function, let's finish with the implementation of the `mutex_lock` function. In the simplest way, the lock will be acquired successfully by a process and the `__mutex_fastpath_lock` will be finished. In this case, we just call the

```
mutex_set_owner(lock);
```

in the end of the `mutex_lock`. The `mutex_set_owner` function is defined in the kernel/locking/mutex.h header file and just sets owner of a lock to the current process:

```
static inline void mutex_set_owner(struct mutex *lock)
{
        lock->owner = current;
}
```

In other way, let's consider situation when a process which wants to acquire a lock is unable to do it, because another process already acquired the same lock. We already know that the `__mutex_lock_slowpath` function will be called in this case. Let's consider implementation of this function. This function is defined in the kernel/locking/mutex.c source code file and starts from the obtaining of the proper mutex by the mutex state given from the `__mutex_fastpath_lock` with the `container_of` macro:

```
__visible void __sched
__mutex_lock_slowpath(atomic_t *lock_count)
```

```
{
        struct mutex *lock = container_of(lock_count, struct mutex, count);

        __mutex_lock_common(lock, TASK_UNINTERRUPTIBLE, 0,
                            NULL, _RET_IP_, NULL, 0);
}
```

and call the `__mutex_lock_common` function with the obtained `mutex`. The `__mutex_lock_common` function starts from preemption disabling until rescheduling:

```
preempt_disable();
```

After this comes the stage of optimistic spinning. As we already know this stage depends on the `CONFIG_MUTEX_SPIN_ON_OWNER` kernel configuration option. If this option is disabled, we skip this stage and move at the last path - `slowpath` of a `mutex` acquisition:

```
if (mutex_optimistic_spin(lock, ww_ctx, use_ww_ctx)) {
        preempt_enable();
        return 0;
}
```

First of all the `mutex_optimistic_spin` function check that we don't need to reschedule or in other words there are no other tasks ready to run that have higher priority. If this check was successful we need to update `MCS` lock wait queue with the current spin. In this way only one spinner can complete for the mutex at one time:

```
osq_lock(&lock->osq)
```

At the next step we start to spin in the next loop:

```
while (true) {
    owner = READ_ONCE(lock->owner);

    if (owner && !mutex_spin_on_owner(lock, owner))
        break;

    if (mutex_try_to_acquire(lock)) {
        lock_acquired(&lock->dep_map, ip);

        mutex_set_owner(lock);
        osq_unlock(&lock->osq);
        return true;
    }
}
```

and try to acquire a lock. First of all we try to take current owner and if the owner exists (it may not exists in a case when a process already released a mutex) and we wait for it in the `mutex_spin_on_owner` function before the owner will release a lock. If new task with higher priority have appeared during wait of the lock owner, we break the loop and go to sleep. In other case, the process already may release a lock, so we try to acquire a lock with the `mutex_try_to_acquired`. If this operation finished successfully, we set new owner for the given mutex, removes ourself from the `MCS` wait queue and exit from the `mutex_optimistic_spin` function. At this state a lock will be acquired by a process and we enable preemption and exit from the `__mutex_lock_common` function:

```
if (mutex_optimistic_spin(lock, ww_ctx, use_ww_ctx)) {
    preempt_enable();
    return 0;
}
```

That's all for this case.

In other case all may not be so successful. For example new task may occur during we spinning in the loop from the `mutex_optimistic_spin` or even we may not get to this loop from the `mutex_optimistic_spin` in a case when there were task(s) with higher priority before this loop. Or finally the `CONFIG_MUTEX_SPIN_ON_OWNER` kernel configuration option disabled. In this case the `mutex_optimistic_spin` will do nothing:

```
#ifndef CONFIG_MUTEX_SPIN_ON_OWNER
static bool mutex_optimistic_spin(struct mutex *lock,
                                  struct ww_acquire_ctx *ww_ctx, const bool use_ww_ctx)
{
    return false;
}
#endif
```

In all of these cases, the `__mutex_lock_common` function will acct like a `semaphore`. We try to acquire a lock again because the owner of a lock might already release a lock before this time:

```
if (!mutex_is_locked(lock) &&
   (atomic_xchg_acquire(&lock->count, 0) == 1))
      goto skip_wait;
```

In a failure case the process which wants to acquire a lock will be added to the waiters list

```
list_add_tail(&waiter.list, &lock->wait_list);
waiter.task = task;
```

In a successful case we update the owner of a lock, enable preemption and exit from the `__mutex_lock_common` function:

```
skip_wait:
        mutex_set_owner(lock);
        preempt_enable();
        return 0;
```

In this case a lock will be acquired. If can't acquire a lock for now, we enter into the following loop:

```
for (;;) {

    if (atomic_read(&lock->count) >= 0 && (atomic_xchg_acquire(&lock->count, -1) == 1))
        break;

    if (unlikely(signal_pending_state(state, task))) {
        ret = -EINTR;
        goto err;
    }

    __set_task_state(task, state);

     schedule_preempt_disabled();
}
```

where try to acquire a lock again and exit if this operation was successful. Yes, we try to acquire a lock again right after unsuccessful try before the loop. We need to do it to make sure that we get a wakeup once a lock will be unlocked. Besides this, it allows us to acquire a lock after sleep. In other case we check the current process for pending signals and exit if the process was interrupted by a `signal` during wait for a lock acquisition. In the end of loop we didn't acquire a lock, so we set the task state for `TASK_UNINTERRUPTIBLE` and go to sleep with call of the `schedule_preempt_disabled` function.

That's all. We have considered all three possible paths through which a process may pass when it will wan to acquire a lock. Now let's consider how `mutex_unlock` is implemented. When the `mutex_unlock` will be called by a process which wants to release a lock, the `__mutex_fastpath_unlock` will be called from the arch/x86/include/asm/mutex_64.h header file:

```
void __sched mutex_unlock(struct mutex *lock)
{
    __mutex_fastpath_unlock(&lock->count, __mutex_unlock_slowpath);
}
```

Implementation of the `__mutex_fastpath_unlock` function is very similar to the implementation of the `__mutex_fastpath_lock` function:

```
static inline void __mutex_fastpath_unlock(atomic_t *v,
                                           void (*fail_fn)(atomic_t *))
{
       asm_volatile_goto(LOCK_PREFIX "   incl %0\n"
                         "   jg %l[exit]\n"
                         : : "m" (v->counter)
                         : "memory", "cc"
                         : exit);
       fail_fn(v);
exit:
       return;
}
```

Actually, there is only one difference. We increment value if the `mutex->count`. So it will represent `unlocked` state after this operation. As `mutex` released, but we have something in the `wait queue` we need to update it. In this case the `fail_fn` function will be called which is `__mutex_unlock_slowpath`. The `__mutex_unlock_slowpath` function just gets the correct `mutex` instance by the given `mutex->count` and calls the `__mutex_unlock_common_slowpath` function:

```
__mutex_unlock_slowpath(atomic_t *lock_count)
{
      struct mutex *lock = container_of(lock_count, struct mutex, count);

      __mutex_unlock_common_slowpath(lock, 1);
}
```

In the `__mutex_unlock_common_slowpath` function we will get the first entry from the wait queue if the wait queue is not empty and wakeup related process:

```
if (!list_empty(&lock->wait_list)) {
    struct mutex_waiter *waiter =
           list_entry(lock->wait_list.next, struct mutex_waiter, list);
                wake_up_process(waiter->task);
}
```

After this, a mutex will be released by previous process and will be acquired by another process from a wait queue.

That's all. We have considered main `API` for manipulation with `mutexes`: `mutex_lock` and `mutex_unlock`. Besides this the Linux kernel provides following API:

- `mutex_lock_interruptible`;
- `mutex_lock_killable`;
- `mutex_trylock`.

and corresponding versions of `unlock` prefixed functions. This part will not describe this `API`, because it is similar to corresponding `API` of `semaphores`. More about it you may read in the previous part.

That's all.

Conclusion

This is the end of the fourth part of the synchronization primitives chapter in the Linux kernel. In this part we met with new synchronization primitive which is called - `mutex`. From the theoretical side, this synchronization primitive very similar on a semaphore. Actually, `mutex` represents binary semaphore. But its implementation differs from the implementation of `semaphore` in the Linux kernel. In the next part we will continue to dive into synchronization primitives in the Linux kernel.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- Mutex
- Spinlock
- Semaphore
- Synchronization primitives
- API
- Locking mechanism
- Context switches
- lock validator
- Atomic
- MCS lock
- Doubly linked list
- x86_64
- Inline assembly
- Memory barrier
- Lock instruction
- JNS instruction
- preemption
- Unix signals
- Previous part

Synchronization primitives in the Linux kernel. Part 5.

Introduction

This is the fifth part of the chapter which describes synchronization primitives in the Linux kernel and in the previous parts we finished to consider different types spinlocks, semaphore and mutex synchronization primitives. We will continue to learn synchronization primitives in this part and start to consider special type of synchronization primitives - readers–writer lock.

The first synchronization primitive of this type will be already familiar for us - semaphore. As in all previous parts of this book, before we will consider implementation of the `reader/writer semaphores` in the Linux kernel, we will start from the theoretical side and will try to understand what is the difference between `reader/writer semaphores` and `normal semaphores`.

So, let's start.

Reader/Writer semaphore

Actually there are two types of operations may be performed on the data. We may read data and make changes in data. Two fundamental operations - `read` and `write`. Usually (but not always), `read` operation is performed more often than `write` operation. In this case, it would be logical to lock data in such a way, that some processes may read locked data in one time, on condition that no one will not change the data. The readers/writer lock allows us to get this lock.

When a process which wants to write something into data, all other `writer` and `reader` processes will be blocked until the process which acquired a lock, will not release it. When a process reads data, other processes which want to read the same data too, will not be locked and will be able to do this. As you may guess, implementation of the `reader/writer semaphore` is based on the implementation of the `normal semaphore`. We already familiar with the semaphore synchronization primitive from the third part of this chapter. From the theoretical side everything looks pretty simple. Let's look how `reader/writer semaphore` is represented in the Linux kernel.

The `semaphore` is represented by the:

```
struct semaphore {
    raw_spinlock_t        lock;
    unsigned int        count;
    struct list_head    wait_list;
};
```

structure. If you will look in the include/linux/rwsem.h header file, you will find definition of the `rw_semaphore` structure which represents `reader/writer semaphore` in the Linux kernel. Let's look at the definition of this structure:

```
#ifdef CONFIG_RWSEM_GENERIC_SPINLOCK
#include <linux/rwsem-spinlock.h>
#else
struct rw_semaphore {
        long count;
        struct list_head wait_list;
        raw_spinlock_t wait_lock;
#ifdef CONFIG_RWSEM_SPIN_ON_OWNER
        struct optimistic_spin_queue osq;
        struct task_struct *owner;
#endif
#ifdef CONFIG_DEBUG_LOCK_ALLOC
        struct lockdep_map      dep_map;
#endif
};
```

Before we will consider fields of the `rw_semaphore` structure, we may notice, that declaration of the `rw_semaphore` structure depends on the `CONFIG_RWSEM_GENERIC_SPINLOCK` kernel configuration option. This option is disabled for the x86_64 architecture by default. We can be sure in this by looking at the corresponding kernel configuration file. In our case, this configuration file is - arch/x86/um/Kconfig:

```
config RWSEM_XCHGADD_ALGORITHM
    def_bool 64BIT

config RWSEM_GENERIC_SPINLOCK
    def_bool !RWSEM_XCHGADD_ALGORITHM
```

So, as this book describes only x86_64 architecture related stuff, we will skip the case when the `CONFIG_RWSEM_GENERIC_SPINLOCK` kernel configuration is enabled and consider definition of the `rw_semaphore` structure only from the include/linux/rwsem.h header file.

If we will take a look at the definition of the `rw_semaphore` structure, we will notice that first three fields are the same that in the `semaphore` structure. It contains `count` field which represents amount of available resources, the `wait_list` field which represents doubly linked list of processes which are waiting to acquire a lock and `wait_lock` spinlock for protection of this list. Notice that `rw_semaphore.count` field is `long` type unlike the same field in the `semaphore` structure.

The `count` field of a `rw_semaphore` structure may have following values:

- `0x0000000000000000` - `reader/writer semaphore` is in unlocked state and no one is waiting for a lock;
- `0x000000000000000X` - `X` readers are active or attempting to acquire a lock and no writer waiting;
- `0xffffffff0000000X` - may represent different cases. The first is - `X` readers are active or attempting to acquire a lock with waiters for the lock. The second is - one writer attempting a lock, no waiters for the lock. And the last - one writer is active and no waiters for the lock;
- `0xffffffff00000001` - may represented two different cases. The first is - one reader is active or attempting to acquire a lock and exist waiters for the lock. The second case is one writer is active or attempting to acquire a lock and no waiters for the lock;
- `0xffffffff00000000` - represents situation when there are readers or writers are queued, but no one is active or is in the process of acquire of a lock;
- `0xfffffffe00000001` - a writer is active or attempting to acquire a lock and waiters are in queue.

So, besides the `count` field, all of these fields are similar to fields of the `semaphore` structure. Last three fields depend on the two configuration options of the Linux kernel: the `CONFIG_RWSEM_SPIN_ON_OWNER` and `CONFIG_DEBUG_LOCK_ALLOC`. The first two fields may be familiar us by declaration of the mutex structure from the previous part. The first `osq` field represents MCS lock spinner for `optimistic spinning` and the second represents process which is current owner of a lock.

The last field of the `rw_semaphore` structure is - `dep_map` - debugging related, and as I already wrote in previous parts, we will skip debugging related stuff in this chapter.

That's all. Now we know a little about what is it `reader/writer lock` in general and `reader/writer semaphore` in particular. Additionally we saw how a `reader/writer semaphore` is represented in the Linux kernel. In this case, we may go ahead and start to look at the API which the Linux kernel provides for manipulation of `reader/writer semaphores`.

Reader/Writer semaphore API

So, we know a little about `reader/writer semaphores` from theoretical side, let's look on its implementation in the Linux kernel. All `reader/writer semaphores` related API is located in the include/linux/rwsem.h header file.

As always Before we will consider an API of the `reader/writer semaphore` mechanism in the Linux kernel, we need to know how to initialize the `rw_semaphore` structure. As we already saw in previous parts of this chapter, all synchronization primitives may be initialized in two ways:

* `statically` ;
* `dynamically` .

And `reader/writer semaphore` is not an exception. First of all, let's take a look at the first approach. We may initialize `rw_semaphore` structure with the help of the `DECLARE_RWSEM` macro in compile time. This macro is defined in the include/linux/rwsem.h header file and looks:

```
#define DECLARE_RWSEM(name) \
        struct rw_semaphore name = __RWSEM_INITIALIZER(name)
```

As we may see, the `DECLARE_RWSEM` macro just expands to the definition of the `rw_semaphore` structure with the given name. Additionally new `rw_semaphore` structure is initialized with the value of the `__RWSEM_INITIALIZER` macro:

```
#define __RWSEM_INITIALIZER(name)              \
{                                                              \
        .count = RWSEM_UNLOCKED_VALUE,                         \
        .wait_list = LIST_HEAD_INIT((name).wait_list),         \
        .wait_lock = __RAW_SPIN_LOCK_UNLOCKED(name.wait_lock)  \
         __RWSEM_OPT_INIT(name)                                \
         __RWSEM_DEP_MAP_INIT(name)
}
```

and expands to the initialization of fields of `rw_semaphore` structure. First of all we initialize `count` field of the `rw_semaphore` structure to the `unlocked` state with `RWSEM_UNLOCKED_VALUE` macro from the arch/x86/include/asm/rwsem.h architecture specific header file:

```
#define RWSEM_UNLOCKED_VALUE            0x00000000L
```

After this we initialize list of a lock waiters with the empty linked list and spinlock for protection of this list with the `unlocked` state too. The `__RWSEM_OPT_INIT` macro depends on the state of the `CONFIG_RWSEM_SPIN_ON_OWNER` kernel configuration option and if this option is enabled it expands to the initialization of the `osq` and `owner` fields of the `rw_semaphore` structure. As we already saw above, the `CONFIG_RWSEM_SPIN_ON_OWNER` kernel configuration option is enabled by default for x86_64 architecture, so let's take a look at the definition of the `__RWSEM_OPT_INIT` macro:

```
#ifdef CONFIG_RWSEM_SPIN_ON_OWNER
    #define __RWSEM_OPT_INIT(lockname) , .osq = OSQ_LOCK_UNLOCKED, .owner = NULL
#else
    #define __RWSEM_OPT_INIT(lockname)
#endif
```

As we may see, the `__RWSEM_OPT_INIT` macro initializes the MCS lock lock with `unlocked` state and initial `owner` of a lock with `NULL` . From this moment, a `rw_semaphore` structure will be initialized in a compile time and may be used for data protection.

The second way to initialize a `rw_semaphore` structure is `dynamically` or use the `init_rwsem` macro from the include/linux/rwsem.h header file. This macro declares an instance of the `lock_class_key` which is related to the lock validator of the Linux kernel and to the call of the `__init_rwsem` function with the given `reader/writer semaphore` :

```
#define init_rwsem(sem)                         \
do {                                                            \
        static struct lock_class_key __key;                     \
                                                                \
        __init_rwsem((sem), #sem, &__key);                      \
} while (0)
```

If you will start definition of the `__init_rwsem` function, you will notice that there are couple of source code files which contain it. As you may guess, sometimes we need to initialize additional fields of the `rw_semaphore` structure, like the `osq` and `owner`. But sometimes not. All of this depends on some kernel configuration options. If we will look at the kernel/locking/Makefile makefile, we will see following lines:

```
obj-$(CONFIG_RWSEM_GENERIC_SPINLOCK) += rwsem-spinlock.o
obj-$(CONFIG_RWSEM_XCHGADD_ALGORITHM) += rwsem-xadd.o
```

As we already know, the Linux kernel for `x86_64` architecture has enabled `CONFIG_RWSEM_XCHGADD_ALGORITHM` kernel configuration option by default:

```
config RWSEM_XCHGADD_ALGORITHM
    def_bool 64BIT
```

in the arch/x86/um/Kconfig kernel configuration file. In this case, implementation of the `__init_rwsem` function will be located in the kernel/locking/rwsem-xadd.c source code file for us. Let's take a look at this function:

```
void __init_rwsem(struct rw_semaphore *sem, const char *name,
                  struct lock_class_key *key)
{
#ifdef CONFIG_DEBUG_LOCK_ALLOC
        debug_check_no_locks_freed((void *)sem, sizeof(*sem));
        lockdep_init_map(&sem->dep_map, name, key, 0);
#endif
        sem->count = RWSEM_UNLOCKED_VALUE;
        raw_spin_lock_init(&sem->wait_lock);
        INIT_LIST_HEAD(&sem->wait_list);
#ifdef CONFIG_RWSEM_SPIN_ON_OWNER
        sem->owner = NULL;
        osq_lock_init(&sem->osq);
#endif
}
```

We may see here almost the same as in `__RWSEM_INITIALIZER` macro with difference that all of this will be executed in runtime.

So, from now we are able to initialize a `reader/writer semaphore` let's look at the `lock` and `unlock` API. The Linux kernel provides following primary API to manipulate `reader/writer semaphores`:

- `void down_read(struct rw_semaphore *sem)` - lock for reading;
- `int down_read_trylock(struct rw_semaphore *sem)` - try lock for reading;
- `void down_write(struct rw_semaphore *sem)` - lock for writing;
- `int down_write_trylock(struct rw_semaphore *sem)` - try lock for writing;
- `void up_read(struct rw_semaphore *sem)` - release a read lock;
- `void up_write(struct rw_semaphore *sem)` - release a write lock;

Let's start as always from the locking. First of all let's consider implementation of the `down_write` function which executes a try of acquiring of a lock for `write`. This function is kernel/locking/rwsem.c source code file and starts from the call of the macro from the include/linux/kernel.h header file:

```
void __sched down_write(struct rw_semaphore *sem)
{
        might_sleep();
        rwsem_acquire(&sem->dep_map, 0, 0, _RET_IP_);

        LOCK_CONTENDED(sem, __down_write_trylock, __down_write);
        rwsem_set_owner(sem);
}
```

We already met the `might_sleep` macro in the previous part. In short words, Implementation of the `might_sleep` macro depends on the `CONFIG_DEBUG_ATOMIC_SLEEP` kernel configuration option and if this option is enabled, this macro just prints a stack trace if it was executed in atomic context. As this macro is mostly for debugging purpose we will skip it and will go ahead. Additionally we will skip the next macro from the `down_read` function - `rwsem_acquire` which is related to the lock validator of the Linux kernel, because this is topic of other part.

The only two things that remained in the `down_write` function is the call of the `LOCK_CONTENDED` macro which is defined in the include/linux/lockdep.h header file and setting of owner of a lock with the `rwsem_set_owner` function which sets owner to currently running process:

```
static inline void rwsem_set_owner(struct rw_semaphore *sem)
{
        sem->owner = current;
}
```

As you already may guess, the `LOCK_CONTENDED` macro does all job for us. Let's look at the implementation of the `LOCK_CONTENDED` macro:

```
#define LOCK_CONTENDED(_lock, try, lock) \
        lock(_lock)
```

As we may see it just calls the `lock` function which is third parameter of the `LOCK_CONTENDED` macro with the given `rw_semaphore`. In our case the third parameter of the `LOCK_CONTENDED` macro is the `__down_write` function which is architecture specific function and located in the arch/x86/include/asm/rwsem.h header file. Let's look at the implementation of the `__down_write` function:

```
static inline void __down_write(struct rw_semaphore *sem)
{
        __down_write_nested(sem, 0);
}
```

which just executes a call of the `__down_write_nested` function from the same source code file. Let's take a look at the implementation of the `__down_write_nested` function:

```
static inline void __down_write_nested(struct rw_semaphore *sem, int subclass)
{
        long tmp;

        asm volatile("# beginning down_write\n\t"
                     LOCK_PREFIX "  xadd      %1,(%2)\n\t"
                     "  test " __ASM_SEL(%w1,%k1) "," __ASM_SEL(%w1,%k1) "\n\t"
                     "  jz        1f\n"
                     "  call call_rwsem_down_write_failed\n"
                     "1:\n"
                     "# ending down_write"
                     : "+m" (sem->count), "=d" (tmp)
                     : "a" (sem), "1" (RWSEM_ACTIVE_WRITE_BIAS)
                     : "memory", "cc");
}
```

As for other synchronization primitives which we saw in this chapter, usually `lock/unlock` functions consists only from an inline assembly statement. As we may see, in our case the same for `__down_write_nested` function. Let's try to understand what does this function do. The first line of our assembly statement is just a comment, let's skip it. The second like contains `LOCK_PREFIX` which will be expanded to the LOCK instruction as we already know. The next xadd instruction executes `add` and `exchange` operations. In other words, `xadd` instruction adds value of the `RWSEM_ACTIVE_WRITE_BIAS` :

```
#define RWSEM_ACTIVE_WRITE_BIAS         (RWSEM_WAITING_BIAS + RWSEM_ACTIVE_BIAS)
```

```
#define RWSEM_WAITING_BIAS              (-RWSEM_ACTIVE_MASK-1)
#define RWSEM_ACTIVE_BIAS               0x00000001L
```

or `0xffffffff00000001` to the `count` of the given `reader/writer semaphore` and returns previous value of it. After this we check the active mask in the `rw_semaphore->count`. If it was zero before, this means that there were no-one writer before, so we acquired a lock. In other way we call the `call_rwsem_down_write_failed` function from the arch/x86/lib/rwsem.S assembly file. The `call_rwsem_down_write_failed` function just calls the `rwsem_down_write_failed` function from the kernel/locking/rwsem-xadd.c source code file anticipatorily save general purpose registers:

```
ENTRY(call_rwsem_down_write_failed)
    FRAME_BEGIN
    save_common_regs
    movq %rax,%rdi
    call rwsem_down_write_failed
    restore_common_regs
    FRAME_END
    ret
    ENDPROC(call_rwsem_down_write_failed)
```

The `rwsem_down_write_failed` function starts from the atomic update of the `count` value:

```
 __visible
struct rw_semaphore __sched *rwsem_down_write_failed(struct rw_semaphore *sem)
{
    count = rwsem_atomic_update(-RWSEM_ACTIVE_WRITE_BIAS, sem);
    ...
    ...
    ...
}
```

with the `-RWSEM_ACTIVE_WRITE_BIAS` value. The `rwsem_atomic_update` function is defined in the arch/x86/include/asm/rwsem.h header file and implement exchange and add logic:

```
static inline long rwsem_atomic_update(long delta, struct rw_semaphore *sem)
{
        return delta + xadd(&sem->count, delta);
}
```

This function atomically adds the given delta to the `count` and returns old value of the count. After this it just returns sum of the given `delta` and old value of the `count` field. In our case we undo write bias from the `count` as we didn't acquire a lock. After this step we try to do `optimistic spinning` by the call of the `rwsem_optimistic_spin` function:

```
if (rwsem_optimistic_spin(sem))
      return sem;
```

We will skip implementation of the `rwsem_optimistic_spin` function, as it is similar on the `mutex_optimistic_spin` function which we saw in the previous part. In short words we check existence other tasks ready to run that have higher priority in the `rwsem_optimistic_spin` function. If there are such tasks, the process will be added to the MCS `waitqueue` and start to spin in the loop until a lock will be able to be acquired. If `optimistic spinning` is disabled, a process will be added to the and marked as waiting for write:

```
waiter.task = current;
waiter.type = RWSEM_WAITING_FOR_WRITE;

if (list_empty(&sem->wait_list))
    waiting = false;
```

```
list_add_tail(&waiter.list, &sem->wait_list);
```

waiters list and start to wait until it will successfully acquire the lock. After we have added a process to the waiters list which was empty before this moment, we update the value of the `rw_semaphore->count` with the `RWSEM_WAITING_BIAS` :

```
count = rwsem_atomic_update(RWSEM_WAITING_BIAS, sem);
```

with this we mark `rw_semaphore->counter` that it is already locked and exists/waits one `writer` which wants to acquire the lock. In other way we try to wake `reader` processes from the `wait queue` that were queued before this `writer` process and there are no active readers. In the end of the `rwsem_down_write_failed` a `writer` process will go to sleep which didn't acquire a lock in the following loop:

```
while (true) {
    if (rwsem_try_write_lock(count, sem))
        break;
    raw_spin_unlock_irq(&sem->wait_lock);
    do {
        schedule();
        set_current_state(TASK_UNINTERRUPTIBLE);
    } while ((count = sem->count) & RWSEM_ACTIVE_MASK);
    raw_spin_lock_irq(&sem->wait_lock);
}
```

I will skip explanation of this loop as we already met similar functional in the previous part.

That's all. From this moment, our `writer` process will acquire or not acquire a lock depends on the value of the `rw_semaphore->count` field. Now if we will look at the implementation of the `down_read` function which executes a try of acquiring of a lock. We will see similar actions which we saw in the `down_write` function. This function calls different debugging and lock validator related functions/macros:

```
void __sched down_read(struct rw_semaphore *sem)
{
        might_sleep();
        rwsem_acquire_read(&sem->dep_map, 0, 0, _RET_IP_);

        LOCK_CONTENDED(sem, __down_read_trylock, __down_read);
}
```

and does all job in the `__down_read` function. The `__down_read` consists of inline assembly statement:

```
static inline void __down_read(struct rw_semaphore *sem)
{
         asm volatile("# beginning down_read\n\t"
                     LOCK_PREFIX _ASM_INC "(%1)\n\t"
                     "  jns        1f\n"
                     "  call call_rwsem_down_read_failed\n"
                     "1:\n\t"
                     "# ending down_read\n\t"
                     : "+m" (sem->count)
                     : "a" (sem)
                     : "memory", "cc");
}
```

which increments value of the given `rw_semaphore->count` and call the `call_rwsem_down_read_failed` if this value is negative. In other way we jump at the label `1:` and exit. After this `read` lock will be successfully acquired. Notice that we check a sign of the `count` value as it may be negative, because as you may remember most significant word of the `rw_semaphore->count` contains negated number of active writers.

Let's consider case when a process wants to acquire a lock for `read` operation, but it is already locked. In this case the `call_rwsem_down_read_failed` function from the arch/x86/lib/rwsem.S assembly file will be called. If you will look at the implementation of this function, you will notice that it does the same that `call_rwsem_down_read_failed` function does. Except it calls the `rwsem_down_read_failed` function instead of `rwsem_dow_write_failed` . Now let's consider implementation of the `rwsem_down_read_failed` function. It starts from the adding a process to the `wait queue` and updating of value of the `rw_semaphore->counter` :

```
long adjustment = -RWSEM_ACTIVE_READ_BIAS;

waiter.task = tsk;
waiter.type = RWSEM_WAITING_FOR_READ;

if (list_empty(&sem->wait_list))
    adjustment += RWSEM_WAITING_BIAS;
list_add_tail(&waiter.list, &sem->wait_list);

count = rwsem_atomic_update(adjustment, sem);
```

Notice that if the `wait queue` was empty before we clear the `rw_semaphore->counter` and undo `read` bias in other way. At the next step we check that there are no active locks and we are first in the `wait queue` we need to join currently active `reader` processes. In other way we go to sleep until a lock will not be able to acquired.

That's all. Now we know how `reader` and `writer` processes will behave in different cases during a lock acquisition. Now let's take a short look at `unlock` operations. The `up_read` and `up_write` functions allows us to unlock a `reader` or `writer` lock. First of all let's take a look at the implementation of the `up_write` function which is defined in the kernel/locking/rwsem.c source code file:

```
void up_write(struct rw_semaphore *sem)
{
        rwsem_release(&sem->dep_map, 1, _RET_IP_);

        rwsem_clear_owner(sem);
        __up_write(sem);
}
```

First of all it calls the `rwsem_release` macro which is related to the lock validator of the Linux kernel, so we will skip it now. And at the next line the `rwsem_clear_owner` function which as you may understand from the name of this function, just clears the `owner` field of the given `rw_semaphore` :

```
static inline void rwsem_clear_owner(struct rw_semaphore *sem)
{
    sem->owner = NULL;
}
```

The `__up_write` function does all job of unlocking of the lock. The `_up_write` is architecture-specific function, so for our case it will be located in the arch/x86/include/asm/rwsem.h source code file. If we will take a look at the implementation of this function, we will see that it does almost the same that `__down_write` function, but conversely. Instead of adding of the `RWSEM_ACTIVE_WRITE_BIAS` to the `count` , we subtract the same value and check the `sign` of the previous value.

If the previous value of the `rw_semaphore->count` is not negative, a writer process released a lock and now it may be acquired by someone else. In other case, the `rw_semaphore->count` will contain negative values. This means that there is at least one `writer` in a wait queue. In this case the `call_rwsem_wake` function will be called. This function acts like similar functions which we already saw above. It store general purpose registers at the stack for preserving and call the `rwsem_wake` function.

First of all the `rwsem_wake` function checks if a spinner is present. In this case it will just acquire a lock which is just released by lock owner. In other case there must be someone in the `wait queue` and we need to wake or writer process if it exists at the top of the `wait queue` or all `reader` processes. The `up_read` function which release a `reader` lock acts in similar way like `up_write` , but with a little difference. Instead of subtracting of `RWSEM_ACTIVE_WRITE_BIAS` from the `rw_semaphore->count` , it subtracts `1` from it, because less significant word of the `count` contains number active locks. After this it checks `sign` of the `count` and calls the `rwsem_wake` like `__up_write` if the `count` is negative or in other way lock will be successfully released.

That's all. We have considered API for manipulation with `reader/writer semaphore` : `up_read/up_write` and `down_read/down_write` . We saw that the Linux kernel provides additional API, besides this functions, like the `,` and etc. But I will not consider implementation of these function in this part because it must be similar on that we have seen in this part of except few subtleties.

Conclusion

This is the end of the fifth part of the synchronization primitives chapter in the Linux kernel. In this part we met with special type of `semaphore` - `readers/writer` semaphore which provides access to data for multiply process to read or for one process to writer. In the next part we will continue to dive into synchronization primitives in the Linux kernel.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- Synchronization primitives
- Readers/Writer lock
- Spinlocks
- Semaphore
- Mutex
- x86_64 architecture
- Doubly linked list
- MCS lock
- API
- Linux kernel lock validator
- Atomic operations
- Inline assembly
- XADD instruction
- LOCK instruction
- Previous part

Synchronization primitives in the Linux kernel. Part 6.

Introduction

This is the sixth part of the chapter which describes synchronization primitives in the Linux kernel and in the previous parts we finished to consider different readers-writer lock synchronization primitives. We will continue to learn synchronization primitives in this part and start to consider a similar synchronization primitive which can be used to avoid the `writer starvation` problem. The name of this synchronization primitive is - `seqlock` or `sequential locks`.

We know from the previous part that readers-writer lock is a special lock mechanism which allows concurrent access for read-only operations, but an exclusive lock is needed for writing or modifying data. As we may guess, it may lead to a problem which is called `writer starvation`. In other words, a writer process can't acquire a lock as long as at least one reader process which acquired a lock holds it. So, in the situation when contention is high, it will lead to situation when a writer process which wants to acquire a lock will wait for it for a long time.

The `seqlock` synchronization primitive can help solve this problem.

As in all previous parts of this book, we will try to consider this synchronization primitive from the theoretical side and only than we will consider API provided by the Linux kernel to manipulate with `seqlocks`.

So, let's start.

Sequential lock

So, what is a `seqlock` synchronization primitive and how does it work? Let's try to answer on these questions in this paragraph. Actually `sequential locks` were introduced in the Linux kernel 2.6.x. Main point of this synchronization primitive is to provide fast and lock-free access to shared resources. Since the heart of `sequential lock` synchronization primitive is spinlock synchronization primitive, `sequential locks` work in situations where the protected resources are small and simple. Additionally write access must be rare and also should be fast.

Work of this synchronization primitive is based on the sequence of events counter. Actually a `sequential lock` allows free access to a resource for readers, but each reader must check existence of conflicts with a writer. This synchronization primitive introduces a special counter. The main algorithm of work of `sequential locks` is simple: Each writer which acquired a sequential lock increments this counter and additionally acquires a spinlock. When this writer finishes, it will release the acquired spinlock to give access to other writers and increment the counter of a sequential lock again.

Read only access works on the following principle, it gets the value of a `sequential lock` counter before it will enter into critical section and compares it with the value of the same `sequential lock` counter at the exit of critical section. If their values are equal, this means that there weren't writers for this period. If their values are not equal, this means that a writer has incremented the counter during the critical section. This conflict means that reading of protected data must be repeated.

That's all. As we may see principle of work of `sequential locks` is simple.

```
unsigned int seq_counter_value;

do {
    seq_counter_value = get_seq_counter_val(&the_lock);
    //
    // do as we want here
    //
} while (__retry__);
```

Actually the Linux kernel does not provide `get_seq_counter_val()` function. Here it is just a stub. Like a `__retry__` too. As I already wrote above, we will see actual the API for this in the next paragraph of this part.

Ok, now we know what a `seqlock` synchronization primitive is and how it is represented in the Linux kernel. In this case, we may go ahead and start to look at the API which the Linux kernel provides for manipulation of synchronization primitives of this type.

Sequential lock API

So, now we know a little about `sequential lock` synchronization primitive from theoretical side, let's look at its implementation in the Linux kernel. All `sequential locks` API are located in the include/linux/seqlock.h header file.

First of all we may see that the a `sequential lock` mechanism is represented by the following type:

```
typedef struct {
    struct seqcount seqcount;
    spinlock_t lock;
} seqlock_t;
```

As we may see the `seqlock_t` provides two fields. These fields represent a sequential lock counter, description of which we saw above and also a spinlock which will protect data from other writers. Note that the `seqcount` counter represented as `seqcount` type. The `seqcount` is structure:

```
typedef struct seqcount {
    unsigned sequence;
#ifdef CONFIG_DEBUG_LOCK_ALLOC
    struct lockdep_map dep_map;
#endif
} seqcount_t;
```

which holds counter of a sequential lock and lock validator related field.

As always in previous parts of this chapter, before we will consider an API of `sequential lock` mechanism in the Linux kernel, we need to know how to initialize an instance of `seqlock_t`.

We saw in the previous parts that often the Linux kernel provides two approaches to execute initialization of the given synchronization primitive. The same situation with the `seqlock_t` structure. These approaches allows to initialize a `seqlock_t` in two following:

- `statically` ;
- `dynamically` .

ways. Let's look at the first approach. We are able to initialize a `seqlock_t` statically with the `DEFINE_SEQLOCK` macro:

```
#define DEFINE_SEQLOCK(x) \
        seqlock_t x = __SEQLOCK_UNLOCKED(x)
```

which is defined in the include/linux/seqlock.h header file. As we may see, the `DEFINE_SEQLOCK` macro takes one argument and expands to the definition and initialization of the `seqlock_t` structure. Initialization occurs with the help of the `__SEQLOCK_UNLOCKED` macro which is defined in the same source code file. Let's look at the implementation of this macro:

```
#define __SEQLOCK_UNLOCKED(lockname)            \
    {                        \
        .seqcount = SEQCNT_ZERO(lockname),     \
        .lock =    __SPIN_LOCK_UNLOCKED(lockname)    \
    }
```

As we may see the, `__SEQLOCK_UNLOCKED` macro executes initialization of fields of the given `seqlock_t` structure. The first field is `seqcount` initialized with the `SEQCNT_ZERO` macro which expands to the:

```
#define SEQCNT_ZERO(lockname) { .sequence = 0, SEQCOUNT_DEP_MAP_INIT(lockname)}
```

So we just initialize counter of the given sequential lock to zero and additionally we can see lock validator related initialization which depends on the state of the `CONFIG_DEBUG_LOCK_ALLOC` kernel configuration option:

```
#ifdef CONFIG_DEBUG_LOCK_ALLOC
# define SEQCOUNT_DEP_MAP_INIT(lockname) \
    .dep_map = { .name = #lockname } \
    ...
    ...
    ...
#else
# define SEQCOUNT_DEP_MAP_INIT(lockname)
    ...
    ...
    ...
#endif
```

As I already wrote in previous parts of this chapter we will not consider debugging and lock validator related stuff in this part. So for now we just skip the `SEQCOUNT_DEP_MAP_INIT` macro. The second field of the given `seqlock_t` is `lock` initialized with the `__SPIN_LOCK_UNLOCKED` macro which is defined in the include/linux/spinlock_types.h header file. We will not consider implementation of this macro here as it just initialize rawspinlock with architecture-specific methods (More abot spinlocks you may read in first parts of this chapter).

We have considered the first way to initialize a sequential lock. Let's consider second way to do the same, but do it dynamically. We can initialize a sequential lock with the `seqlock_init` macro which is defined in the same include/linux/seqlock.h header file.

Let's look at the implementation of this macro:

```
#define seqlock_init(x)                    \
    do {                        \
        seqcount_init(&(x)->seqcount);        \
        spin_lock_init(&(x)->lock);        \
    } while (0)
```

As we may see, the `seqlock_init` expands into two macros. The first macro `seqcount_init` takes counter of the given sequential lock and expands to the call of the `__seqcount_init` function:

```
# define seqcount_init(s)                \
    do {                        \
        static struct lock_class_key __key;    \
        __seqcount_init((s), #s, &__key);    \
    } while (0)
```

from the same header file. This function

```
static inline void __seqcount_init(seqcount_t *s, const char *name,
                      struct lock_class_key *key)
{
    lockdep_init_map(&s->dep_map, name, key, 0);
    s->sequence = 0;
}
```

just initializes counter of the given `seqcount_t` with zero. The second call from the `seqlock_init` macro is the call of the `spin_lock_init` macro which we saw in the first part of this chapter.

So, now we know how to initialize a `sequential lock`, now let's look at how to use it. The Linux kernel provides following API to manipulate `sequential locks`:

```
static inline unsigned read_seqbegin(const seqlock_t *sl);
static inline unsigned read_seqretry(const seqlock_t *sl, unsigned start);
static inline void write_seqlock(seqlock_t *sl);
static inline void write_sequnlock(seqlock_t *sl);
static inline void write_seqlock_irq(seqlock_t *sl);
static inline void write_sequnlock_irq(seqlock_t *sl);
static inline void read_seqlock_excl(seqlock_t *sl)
static inline void read_sequnlock_excl(seqlock_t *sl)
```

and others. Before we move on to considering the implementation of this API, we must know that actually there are two types of readers. The first type of reader never blocks a writer process. In this case writer will not wait for readers. The second type of reader which can lock. In this case, the locking reader will block the writer as it will wait while reader will not release its lock.

First of all let's consider the first type of readers. The `read_seqbegin` function begins a seq-read critical section.

As we may see this function just returns value of the `read_seqcount_begin` function:

```
static inline unsigned read_seqbegin(const seqlock_t *sl)
{
    return read_seqcount_begin(&sl->seqcount);
}
```

In its turn the `read_seqcount_begin` function calls the `raw_read_seqcount_begin` function:

```
static inline unsigned read_seqcount_begin(const seqcount_t *s)
{
    return raw_read_seqcount_begin(s);
}
```

which just returns value of the `sequential lock` counter:

```
static inline unsigned raw_read_seqcount(const seqcount_t *s)
{
    unsigned ret = READ_ONCE(s->sequence);
    smp_rmb();
    return ret;
}
```

After we have the initial value of the given `sequential lock` counter and did some stuff, we know from the previous paragraph of this function, that we need to compare it with the current value of the counter the same `sequential lock` before we will exit from the critical section. We can achieve this by the call of the `read_seqretry` function. This function takes a `sequential lock`, start value of the counter and through a chain of functions:

```
static inline unsigned read_seqretry(const seqlock_t *sl, unsigned start)
{
    return read_seqcount_retry(&sl->seqcount, start);
}

static inline int read_seqcount_retry(const seqcount_t *s, unsigned start)
{
    smp_rmb();
    return __read_seqcount_retry(s, start);
}
```

it calls the `__read_seqcount_retry` function:

```
static inline int __read_seqcount_retry(const seqcount_t *s, unsigned start)
{
    return unlikely(s->sequence != start);
}
```

which just compares value of the counter of the given `sequential lock` with the initial value of this counter. If the initial value of the counter which is obtained from `read_seqbegin()` function is odd, this means that a writer was in the middle of updating the data when our reader began to act. In this case the value of the data can be in inconsistent state, so we need to try to read it again.

This is a common pattern in the Linux kernel. For example, you may remember the `jiffies` concept from the first part of the timers and time management in the Linux kernel chapter. The sequential lock is used to obtain value of `jiffies` at x86_64 architecture:

```
u64 get_jiffies_64(void)
{
    unsigned long seq;
    u64 ret;

    do {
        seq = read_seqbegin(&jiffies_lock);
        ret = jiffies_64;
    } while (read_seqretry(&jiffies_lock, seq));
    return ret;
}
```

Here we just read the value of the counter of the `jiffies_lock` sequential lock and then we write value of the `jiffies_64` system variable to the `ret`. As here we may see `do/while` loop, the body of the loop will be executed at least one time. So, as the body of loop was executed, we read and compare the current value of the counter of the `jiffies_lock` with the initial value. If these values are not equal, execution of the loop will be repeated, else `get_jiffies_64` will return its value in `ret`.

We just saw the first type of readers which do not block writer and other readers. Let's consider second type. It does not update value of a `sequential lock` counter, but just locks `spinlock`:

```
static inline void read_seqlock_excl(seqlock_t *sl)
{
    spin_lock(&sl->lock);
}
```

So, no one reader or writer can't access protected data. When a reader finishes, the lock must be unlocked with the:

```
static inline void read_sequnlock_excl(seqlock_t *sl)
{
    spin_unlock(&sl->lock);
}
```

function.

Now we know how `sequential lock` work for readers. Let's consider how does writer act when it wants to acquire a `sequential lock` to modify data. To acquire a `sequential lock`, writer should use `write_seqlock` function. If we look at the implementation of this function:

```
static inline void write_seqlock(seqlock_t *sl)
{
    spin_lock(&sl->lock);
    write_seqcount_begin(&sl->seqcount);
```

```
}
```

We will see that it acquires `spinlock` to prevent access from other writers and calls the `write_seqcount_begin` function. This function just increments value of the `sequential lock` counter:

```
static inline void raw_write_seqcount_begin(seqcount_t *s)
{
    s->sequence++;
    smp_wmb();
}
```

When a writer process will finish to modify data, the `write_sequnlock` function must be called to release a lock and give access to other writers or readers. Let's consider at the implementation of the `write_sequnlock` function. It looks pretty simple:

```
static inline void write_sequnlock(seqlock_t *sl)
{
    write_seqcount_end(&sl->seqcount);
    spin_unlock(&sl->lock);
}
```

First of all it just calls `write_seqcount_end` function to increase value of the counter of the `sequential` lock again:

```
static inline void raw_write_seqcount_end(seqcount_t *s)
{
    smp_wmb();
    s->sequence++;
}
```

and in the end we just call the `spin_unlock` macro to give access for other readers or writers.

That's all about `sequential lock` mechanism in the Linux kernel. Of course we did not consider full API of this mechanism in this part. But all other functions are based on these which we described here. For example, Linux kernel also provides some safe macros/functions to use `sequential lock` mechanism in interrupt handlers of softirq: `write_seqclock_irq` and `write_sequnlock_irq` :

```
static inline void write_seqlock_irq(seqlock_t *sl)
{
    spin_lock_irq(&sl->lock);
    write_seqcount_begin(&sl->seqcount);
}

static inline void write_sequnlock_irq(seqlock_t *sl)
{
    write_seqcount_end(&sl->seqcount);
    spin_unlock_irq(&sl->lock);
}
```

As we may see, these functions differ only in the initialization of spinlock. They call `spin_lock_irq` and `spin_unlock_irq` instead of `spin_lock` and `spin_unlock` .

Or for example `write_seqlock_irqsave` and `write_sequnlock_irqrestore` functions which are the same but used `spin_lock_irqsave` and `spin_unlock_irqsave` macro to use in IRQ) handlers.

That's all.

Conclusion

This is the end of the sixth part of the synchronization primitives chapter in the Linux kernel. In this part we met with new synchronization primitive which is called - `sequential lock`. From the theoretical side, this synchronization primitive very similar on a readers-writer lock synchronization primitive, but allows to avoid `writer-starving` issue.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides.

Links

- synchronization primitives)
- readers-writer lock
- spinlock
- critical section
- lock validator
- debugging
- API
- x86_64
- Timers and time management in the Linux kernel
- interrupt handlers
- softirq
- IRQ)
- Previous part

Linux kernel memory management

This chapter describes memory management in the linux kernel. You will see here a couple of posts which describe different parts of the linux memory management framework:

- Memblock - describes early `memblock` allocator.
- Fix-Mapped Addresses and ioremap - describes `fix-mapped` addresses and early `ioremap` .
- kmemcheck - third part describes `kmemcheck` tool.

Linux kernel memory management Part 1.

Introduction

Memory management is one of the most complex (and I think that it is the most complex) part of the operating system kernel. In the last preparations before the kernel entry point part we stopped right before call of the `start_kernel` function. This function initializes all the kernel features (including architecture-dependent features) before the kernel runs the first `init` process. You may remember as we built early page tables, identity page tables and fixmap page tables in the boot time. No complicated memory management is working yet. When the `start_kernel` function is called we will see the transition to more complex data structures and techniques for memory management. For a good understanding of the initialization process in the linux kernel we need to have a clear understanding of these techniques. This chapter will provide an overview of the different parts of the linux kernel memory management framework and its API, starting from the `memblock`.

Memblock

Memblock is one of the methods of managing memory regions during the early bootstrap period while the usual kernel memory allocators are not up and running yet. Previously it was called `Logical Memory Block`, but with the patch by Yinghai Lu, it was renamed to the `memblock`. As Linux kernel for `x86_64` architecture uses this method. We already met `memblock` in the Last preparations before the kernel entry point part. And now it's time to get acquainted with it closer. We will see how it is implemented.

We will start to learn `memblock` from the data structures. Definitions of all logical-memory-block-related data structures can be found in the include/linux/memblock.h header file.

The first structure has the same name as this part and it is:

```
struct memblock {
         bool bottom_up;
         phys_addr_t current_limit;
         struct memblock_type memory;   --> array of memblock_region
         struct memblock_type reserved; --> array of memblock_region
#ifdef CONFIG_HAVE_MEMBLOCK_PHYS_MAP
         struct memblock_type physmem;
#endif
};
```

This structure contains five fields. First is `bottom_up` which allows allocating memory in bottom-up mode when it is `true`. Next field is `current_limit`. This field describes the limit size of the memory block. The next three fields describe the type of the memory block. It can be: reserved, memory and physical memory (physical memory is available if the `CONFIG_HAVE_MEMBLOCK_PHYS_MAP` configuration option is enabled). Now we see yet another data structure - `memblock_type`. Let's look at its definition:

```
struct memblock_type {
    unsigned long cnt;
    unsigned long max;
    phys_addr_t total_size;
    struct memblock_region *regions;
};
```

This structure provides information about the memory type. It contains fields which describe the number of memory regions inside the current memory block, the size of all memory regions, the size of the allocated array of the memory regions, and a pointer to the array of the `memblock_region` structures. `memblock_region` is a structure which describes a memory region. Its definition is:

```
struct memblock_region {
        phys_addr_t base;
        phys_addr_t size;
        unsigned long flags;
#ifdef CONFIG_HAVE_MEMBLOCK_NODE_MAP
        int nid;
#endif
};
```

`memblock_region` provides the base address and size of the memory region as well as a flags field which can have the following values:

```
enum {
    MEMBLOCK_NONE    = 0x0,    /* No special request */
    MEMBLOCK_HOTPLUG    = 0x1,    /* hotpluggable region */
    MEMBLOCK_MIRROR    = 0x2,    /* mirrored region */
    MEMBLOCK_NOMAP    = 0x4,    /* don't add to kernel direct mapping */
};
```

Also `memblock_region` provides an integer field - numa node selector, if the `CONFIG_HAVE_MEMBLOCK_NODE_MAP` configuration option is enabled.

Schematically we can imagine it as:

```
+---------------------------+   +---------------------------+
|         memblock          |   |                           |
|  _______________________  |   |                           |
| |        memory         | |   |       Array of the        |
| |      memblock_type    |-|-->|      memblock_region      |
| |_______________________| |   |                           |
|                           |   +---------------------------+
|  _______________________  |   +---------------------------+
| |       reserved        | |   |                           |
| |      memblock_type    |-|-->|       Array of the        |
| |_______________________| |   |      memblock_region      |
|                           |   |                           |
+---------------------------+   +---------------------------+
```

These three structures: `memblock` , `memblock_type` and `memblock_region` are main in the `Memblock` . Now we know about it and can look at Memblock initialization process.

Memblock initialization

As all API of the `memblock` are described in the include/linux/memblock.h header file, all implementations of these functions are in the mm/memblock.c source code file. Let's look at the top of the source code file and we will see the initialization of the `memblock` structure:

```
struct memblock memblock __initdata_memblock = {
    .memory.regions        = memblock_memory_init_regions,
    .memory.cnt            = 1,
    .memory.max            = INIT_MEMBLOCK_REGIONS,

    .reserved.regions    = memblock_reserved_init_regions,
    .reserved.cnt        = 1,
    .reserved.max        = INIT_MEMBLOCK_REGIONS,

#ifdef CONFIG_HAVE_MEMBLOCK_PHYS_MAP
```

```
    .physmem.regions    = memblock_physmem_init_regions,
    .physmem.cnt        = 1,
    .physmem.max        = INIT_PHYSMEM_REGIONS,
#endif
    .bottom_up            = false,
    .current_limit        = MEMBLOCK_ALLOC_ANYWHERE,
};
```

Here we can see initialization of the `memblock` structure which has the same name as structure - `memblock`. First of all note the `__initdata_memblock`. Definition of this macro looks like:

```
#ifdef CONFIG_ARCH_DISCARD_MEMBLOCK
    #define __init_memblock __meminit
    #define __initdata_memblock __meminitdata
#else
    #define __init_memblock
    #define __initdata_memblock
#endif
```

You can see that it depends on `CONFIG_ARCH_DISCARD_MEMBLOCK`. If this configuration option is enabled, memblock code will be put into the `.init` section and will be released after the kernel is booted up.

Next we can see the initialization of the `memblock_type memory`, `memblock_type reserved` and `memblock_type physmem` fields of the `memblock` structure. Here we are interested only in the `memblock_type.regions` initialization process. Note that every `memblock_type` field is initialized by and array of `memblock_region`s:

```
static struct memblock_region memblock_memory_init_regions[INIT_MEMBLOCK_REGIONS] __initdata_memblock;
static struct memblock_region memblock_reserved_init_regions[INIT_MEMBLOCK_REGIONS] __initdata_memblock;
#ifdef CONFIG_HAVE_MEMBLOCK_PHYS_MAP
static struct memblock_region memblock_physmem_init_regions[INIT_PHYSMEM_REGIONS] __initdata_memblock;
#endif
```

Every array contains 128 memory regions. We can see it in the `INIT_MEMBLOCK_REGIONS` macro definition:

```
#define INIT_MEMBLOCK_REGIONS   128
```

Note that all arrays are also defined with the `__initdata_memblock` macro which we already saw in the `memblock` structure initialization (read above if you've forgotten).

The last two fields describe that `bottom_up` allocation is disabled and the limit of the current Memblock is:

```
#define MEMBLOCK_ALLOC_ANYWHERE (~(phys_addr_t)0)
```

which is `0xffffffffffffffff`.

On this step the initialization of the `memblock` structure has been finished and we can have a look at the Memblock API.

Memblock API

Ok we have finished with the initialization of the `memblock` structure and now we can look at the Memblock API and its implementation. As I said above, the implementation of `memblock` is taking place fully in mm/memblock.c. To understand how `memblock` works and how it is implemented, let's look at its usage first. There are a couple of places in the linux kernel where memblock is used. For example let's take `memblock_x86_fill` function from the arch/x86/kernel/e820.c. This function goes through the memory map provided by the e820 and adds memory regions reserved by the kernel to the `memblock` with the `memblock_add` function. Since we have met the `memblock_add` function first, let's start from it.

This function takes a physical base address and the size of the memory region as arguments and add them to the `memblock`. The `memblock_add` function does not do anything special in its body, but just calls the:

```
memblock_add_range(&memblock.memory, base, size, MAX_NUMNODES, 0);
```

function. We pass the memory block type - `memory`, the physical base address and the size of the memory region, the maximum number of nodes which is 1 if `CONFIG_NODES_SHIFT` is not set in the configuration file or `1 << CONFIG_NODES_SHIFT` if it is set, and the flags. The `memblock_add_range` function adds a new memory region to the memory block. It starts by checking the size of the given region and if it is zero it just returns. After this, `memblock_add_range` checks the existence of the memory regions in the `memblock` structure with the given `memblock_type`. If there are no memory regions, we just fill a new `memory_region` with the given values and return (we already saw the implementation of this in the First touch of the linux kernel memory manager framework). If `memblock_type` is not empty, we start to add a new memory region to the `memblock` with the given `memblock_type`.

First of all we get the end of the memory region with the:

```
phys_addr_t end = base + memblock_cap_size(base, &size);
```

`memblock_cap_size` adjusts `size` that `base + size` will not overflow. Its implementation is pretty easy:

```
static inline phys_addr_t memblock_cap_size(phys_addr_t base, phys_addr_t *size)
{
    return *size = min(*size, (phys_addr_t)ULLONG_MAX - base);
}
```

`memblock_cap_size` returns the new size which is the smallest value between the given size and `ULLONG_MAX - base`.

After that we have the end address of the new memory region, `memblock_add_range` checks for overlap and merge conditions with memory regions that have been added before. Insertion of the new memory region to the `memblock` consists of two steps:

- Adding of non-overlapping parts of the new memory area as separate regions;
- Merging of all neighboring regions.

We are going through all the already stored memory regions and checking for overlap with the new region:

```
    for (i = 0; i < type->cnt; i++) {
        struct memblock_region *rgn = &type->regions[i];
        phys_addr_t rbase = rgn->base;
        phys_addr_t rend = rbase + rgn->size;

        if (rbase >= end)
            break;
        if (rend <= base)
            continue;
        ...
        ...
        ...
    }
```

If the new memory region does not overlap with regions which are already stored in the `memblock`, insert this region into the memblock with and this is first step, we check if the new region can fit into the memory block and call `memblock_double_array` in another way:

```
while (type->cnt + nr_new > type->max)
    if (memblock_double_array(type, obase, size) < 0)
        return -ENOMEM;
    insert = true;
    goto repeat;
```

`memblock_double_array` doubles the size of the given regions array. Then we set `insert` to `true` and go to the `repeat` label. In the second step, starting from the `repeat` label we go through the same loop and insert the current memory region into the memory block with the `memblock_insert_region` function:

```
    if (base < end) {
        nr_new++;
        if (insert)
            memblock_insert_region(type, i, base, end - base,
                           nid, flags);
    }
```

Since we set `insert` to `true` in the first step, now `memblock_insert_region` will be called. `memblock_insert_region` has almost the same implementation that we saw when we inserted a new region to the empty `memblock_type` (see above). This function gets the last memory region:

```
struct memblock_region *rgn = &type->regions[idx];
```

and copies the memory area with `memmove` :

```
memmove(rgn + 1, rgn, (type->cnt - idx) * sizeof(*rgn));
```

After this fills `memblock_region` fields of the new memory region base, size, etc. and increases size of the `memblock_type` . In the end of the execution, `memblock_add_range` calls `memblock_merge_regions` which merges neighboring compatible regions in the second step.

In the second case the new memory region can overlap already stored regions. For example we already have `region1` in the `memblock` :

```
0                    0x1000
+-----------------------+
|                       |
|                       |
|        region1        |
|                       |
|                       |
+-----------------------+
```

And now we want to add `region2` to the `memblock` with the following base address and size:

```
0x100                 0x2000
+-----------------------+
|                       |
|                       |
|        region2        |
|                       |
|                       |
+-----------------------+
```

In this case set the base address of the new memory region as the end address of the overlapped region with:

```
base = min(rend, end);
```

So it will be `0x1000` in our case. And insert it as we did it already in the second step with:

```
if (base < end) {
```

```
    nr_new++;
    if (insert)
        memblock_insert_region(type, i, base, end - base, nid, flags);
}
```

In this case we insert `overlapping portion` (we insert only the higher portion, because the lower portion is already in the overlapped memory region), then the remaining portion and merge these portions with `memblock_merge_regions`. As I said above `memblock_merge_regions` function merges neighboring compatible regions. It goes through all memory regions from the given `memblock_type`, takes two neighboring memory regions - `type->regions[i]` and `type->regions[i + 1]` and checks that these regions have the same flags, belong to the same node and that the end address of the first regions is not equal to the base address of the second region:

```
while (i < type->cnt - 1) {
    struct memblock_region *this = &type->regions[i];
    struct memblock_region *next = &type->regions[i + 1];
    if (this->base + this->size != next->base ||
        memblock_get_region_node(this) !=
        memblock_get_region_node(next) ||
        this->flags != next->flags) {
        BUG_ON(this->base + this->size > next->base);
        i++;
        continue;
    }
```

If none of these conditions are true, we update the size of the first region with the size of the next region:

```
this->size += next->size;
```

As we update the size of the first memory region with the size of the next memory region, we move all memory regions which are after the (`next`) memory region one index backwards with the `memmove` function:

```
memmove(next, next + 1, (type->cnt - (i + 2)) * sizeof(*next));
```

The `memmove` here moves all regions which are located after the `next` region to the base address of the `next` region. In the end we just decrease the count of the memory regions which belong to the `memblock_type`:

```
type->cnt--;
```

After this we will get two memory regions merged into one:

```
0                                             0x2000
+-------------------------------------------------+
|                                                 |
|                                                 |
|                   region1                       |
|                                                 |
|                                                 |
+-------------------------------------------------+
```

As we decreased counts of regions in a memblock with certain type, increased size of the `this` region and shifted all regions which are located after `next` region to its place.

That's all. This is the whole principle of the work of the `memblock_add_range` function.

There is also `memblock_reserve` function which does the same as `memblock_add`, but with one difference. It stores `memblock_type.reserved` in the memblock instead of `memblock_type.memory`.

Of course this is not the full API. Memblock provides APIs not only for adding `memory` and `reserved` memory regions, but also:

- memblock_remove - removes memory region from memblock;
- memblock_find_in_range - finds free area in given range;
- memblock_free - releases memory region in memblock;
- for_each_mem_range - iterates through memblock areas.

and many more....

Getting info about memory regions

Memblock also provides an API for getting information about allocated memory regions in the `memblock`. It is split in two parts:

- get_allocated_memblock_memory_regions_info - getting info about memory regions;
- get_allocated_memblock_reserved_regions_info - getting info about reserved regions.

Implementation of these functions is easy. Let's look at `get_allocated_memblock_reserved_regions_info` for example:

```
phys_addr_t __init_memblock get_allocated_memblock_reserved_regions_info(
                    phys_addr_t *addr)
{
    if (memblock.reserved.regions == memblock_reserved_init_regions)
        return 0;

    *addr = __pa(memblock.reserved.regions);

    return PAGE_ALIGN(sizeof(struct memblock_region) *
              memblock.reserved.max);
}
```

First of all this function checks that `memblock` contains reserved memory regions. If `memblock` does not contain reserved memory regions we just return zero. Otherwise we write the physical address of the reserved memory regions array to the given address and return aligned size of the allocated array. Note that there is `PAGE_ALIGN` macro used for align. Actually it depends on size of page:

```
#define PAGE_ALIGN(addr) ALIGN(addr, PAGE_SIZE)
```

Implementation of the `get_allocated_memblock_memory_regions_info` function is the same. It has only one difference, `memblock_type.memory` used instead of `memblock_type.reserved`.

Memblock debugging

There are many calls to `memblock_dbg` in the memblock implementation. If you pass the `memblock=debug` option to the kernel command line, this function will be called. Actually `memblock_dbg` is just a macro which expands to `printk`:

```
#define memblock_dbg(fmt, ...) \
         if (memblock_debug) printk(KERN_INFO pr_fmt(fmt), ##__VA_ARGS__)
```

For example you can see a call of this macro in the `memblock_reserve` function:

```
memblock_dbg("memblock_reserve: [%#016llx-%#016llx] flags %#02lx %pF\n",
             (unsigned long long)base,
             (unsigned long long)base + size - 1,
             flags, (void *)_RET_IP_);
```

And you will see something like this:

```
Kernel command line: root=/dev/sdb earlyprintk=ttyS0 loglevel=7 debug rdinit=/sbin/init root=/dev/ram memblock=debug
memblock_virt_alloc_try_nid_nopanic: 32768 bytes align=0x0 nid=-1 from=0x0 max_addr=0x0 alloc_large_system_hash+0x144/0x228
memblock_reserve: [0x0000023ff38e00-0x0000023ff40dff] flags 0x0 memblock_virt_alloc_internal+0xfd/0x13f
PID hash table entries: 4096 (order: 3, 32768 bytes)
memblock_virt_alloc_try_nid_nopanic: 67108864 bytes align=0x1000 nid=-1 from=0x0 max_addr=0xffffffff swiotlb_init+0x4c/0xad
memblock_reserve: [0x000000bbfe0000-0x000000bffdffff] flags 0x0 memblock_virt_alloc_internal+0xfd/0x13f
memblock_virt_alloc_try_nid_nopanic: 32768 bytes align=0x1000 nid=-1 from=0x0 max_addr=0xffffffff swiotlb_init_with_tbl+0x69/0x147
memblock_reserve: [0x000000bbfd8000-0x000000bbfdffff] flags 0x0 memblock_virt_alloc_internal+0xfd/0x13f
memblock_virt_alloc_try_nid: 131072 bytes align=0x1000 nid=-1 from=0x0 max_addr=0x0 swiotlb_init_with_tbl+0xb9/0x147
memblock_reserve: [0x0000023ff18000-0x0000023ff37fff] flags 0x0 memblock_virt_alloc_internal+0xfd/0x13f
memblock_virt_alloc_try_nid: 262144 bytes align=0x1000 nid=-1 from=0x0 max_addr=0x0 swiotlb_init_with_tbl+0xe8/0x147
memblock_reserve: [0x0000023fed8000-0x0000023ff17fff] flags 0x0 memblock_virt_alloc_internal+0xfd/0x13f
```

Memblock also has support in debugfs. If you run the kernel on another architecture than `X86` you can access:

- /sys/kernel/debug/memblock/memory
- /sys/kernel/debug/memblock/reserved
- /sys/kernel/debug/memblock/physmem

to get a dump of the `memblock` contents.

Conclusion

This is the end of the first part about linux kernel memory management. If you have questions or suggestions, ping me on twitter 0xAX, drop me an email or just create an issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me a PR to linux-insides.

Links

- e820
- numa
- debugfs
- First touch of the linux kernel memory manager framework

Linux kernel memory management Part 2.

Fix-Mapped Addresses and ioremap

`Fix-Mapped` addresses are a set of special compile-time addresses whose corresponding physical addresses do not have to be a linear address minus `__START_KERNEL_map`. Each fix-mapped address maps one page frame and the kernel uses them as pointers that never change their address. That is the main point of these addresses. As the comment says: `to have a constant address at compile time, but to set the physical address only in the boot process`. You can remember that in the earliest part, we already set the `level2_fixmap_pgt`:

```
NEXT_PAGE(level2_fixmap_pgt)
    .fill    506,8,0
    .quad    level1_fixmap_pgt - __START_KERNEL_map + _PAGE_TABLE
    .fill    5,8,0

NEXT_PAGE(level1_fixmap_pgt)
    .fill    512,8,0
```

As you can see `level2_fixmap_pgt` is right after the `level2_kernel_pgt` which is kernel code+data+bss. Every fix-mapped address is represented by an integer index which is defined in the `fixed_addresses` enum from the arch/x86/include/asm/fixmap.h. For example it contains entries for `VSYSCALL_PAGE` - if emulation of legacy vsyscall page is enabled, `FIX_APIC_BASE` for local apic, etc. In virtual memory fix-mapped area is placed in the modules area:

```
         +-----------+-----------------+---------------+------------------+
         |           |                 |               |                  |
         |kernel text|      kernel     |               |    vsyscalls     |
         | mapping   |       text      |    Modules    |    fix-mapped    |
         |from phys 0|       data      |               |    addresses     |
         |           |                 |               |                  |
         +-----------+-----------------+---------------+------------------+
__START_KERNEL_map   __START_KERNEL    MODULES_VADDR            0xffffffffffffffff
```

Base virtual address and size of the `fix-mapped` area are presented by the two following macro:

```
#define FIXADDR_SIZE    (__end_of_permanent_fixed_addresses << PAGE_SHIFT)
#define FIXADDR_START    (FIXADDR_TOP - FIXADDR_SIZE)
```

Here `__end_of_permanent_fixed_addresses` is an element of the `fixed_addresses` enum and as I wrote above: Every fix-mapped address is represented by an integer index which is defined in the `fixed_addresses`. `PAGE_SHIFT` determines the size of a page. For example size of the one page we can get with the `1 << PAGE_SHIFT` expression.

In our case we need to get the size of the fix-mapped area, but not only of one page, that's why we are using `__end_of_permanent_fixed_addresses` for getting the size of the fix-mapped area. The `__end_of_permanent_fixed_addresses` is the last index of the `fixed_addresses` enum or in other words the `__end_of_permanent_fixed_addresses` contains amount of pages in a fixed-mapped area. So if multiply value of the `__end_of_permanent_fixed_addresses` on a page size value we will get size of fix-mapped area. In my case it's a little more than `536` kilobytes. In your case it might be a different number, because the size depends on amount of the fix-mapped addresses which are depends on your kernel's configuration.

The second `FIXADDR_START` macro just subtracts the fix-mapped area size from the last address of the fix-mapped area to get its base virtual address. `FIXADDR_TOP` is a rounded up address from the base address of the vsyscall space:

```
#define FIXADDR_TOP     (round_up(VSYSCALL_ADDR + PAGE_SIZE, 1<<PMD_SHIFT) - PAGE_SIZE)
```

The `fixed_addresses` enums are used as an index to get the virtual address by the `fix_to_virt` function. Implementation of this function is easy:

```
static __always_inline unsigned long fix_to_virt(const unsigned int idx)
{
        BUILD_BUG_ON(idx >= __end_of_fixed_addresses);
        return __fix_to_virt(idx);
}
```

first of all it checks that the index given for the `fixed_addresses` enum is not greater or equal than `__end_of_fixed_addresses` with the `BUILD_BUG_ON` macro and then returns the result of the `__fix_to_virt` macro:

```
#define __fix_to_virt(x)        (FIXADDR_TOP - ((x) << PAGE_SHIFT))
```

Here we shift left the given index of a `fix-mapped` area on the `PAGE_SHIFT` which determines size of a page as I wrote above and subtract it from the `FIXADDR_TOP` which is the highest address of the `fix-mapped` area:

```
+-----------------+
|    PAGE 1       | FIXADDR_TOP (virt address)
|    PAGE 2       |
|    PAGE 3       |
|    PAGE 4 (idx) | x - 4
|    PAGE 5       |
+-----------------+
```

There is an inverse function for getting an index of a fix-mapped area corresponding to the given virtual address:

```
static inline unsigned long virt_to_fix(const unsigned long vaddr)
{
        BUG_ON(vaddr >= FIXADDR_TOP || vaddr < FIXADDR_START);
        return __virt_to_fix(vaddr);
}
```

The `virt_to_fix` takes a virtual address, checks that this address is between `FIXADDR_START` and `FIXADDR_TOP` and calls the `__virt_to_fix` macro which implemented as:

```
#define __virt_to_fix(x)        ((FIXADDR_TOP - ((x)&PAGE_MASK)) >> PAGE_SHIFT)
```

As we may see, the `__virt_to_fix` macro clears the first `12` bits in the given virtual address, subtracts it from the last address the of `fix-mapped` area (`FIXADDR_TOP`) and shifts the result right on `PAGE_SHIFT` which is `12` . Let me explain how it works.

As in previous example (in `__fix_to_virt` macro), we start from the top of the fix-mapped area. We also go back to bottom from the top to search an index of a fix-mapped area corresponding to the given virtual address. As you may see, first of all we will clear the first `12` bits in the given virtual address with `x & PAGE_MASK` expression. This allows us to get base address of page. We need to do this for case when the given virtual address points somewhere in a beginning/middle or end of a page, but not to the base address of it. At the next step subtract this from the `FIXADDR_TOP` and this gives us virtual address of a corresponding page in a fix-mapped area. In the end we just divide value of this address on `PAGE_SHIFT` . This gives us index of a fix-mapped area corresponding to the given virtual address. It may looks hard, but if you will go through this step by step, you will be sure that the `__virt_to_fix` macro is pretty easy.

That's all. For this moment we know a little about `fix-mapped` addresses, but this is enough to go next.

`Fix-mapped` addresses are used in different places in the linux kernel. `IDT` descriptor stored there, Intel Trusted Execution Technology UUID stored in the `fix-mapped` area started from `FIX_TBOOT_BASE` index, Xen bootmap and many more... We already saw a little about `fix-mapped` addresses in the fifth part about of the linux kernel initialization. We use `fix-mapped` area in the early `ioremap` initialization. Let's look at it more closely and try to understand what `ioremap` is, how it is implemented in the kernel and how it is related to the `fix-mapped` addresses.

ioremap

The Linux kernel provides many different primitives to manage memory. For this moment we will touch `I/O memory`. Every device is controlled by reading/writing from/to its registers. For example a driver can turn off/on a device by writing to its registers or get the state of a device by reading from its registers. Besides registers, many devices have buffers where a driver can write something or read from there. As we know for this moment there are two ways to access device's registers and data buffers:

- through the I/O ports;
- mapping of the all registers to the memory address space;

In the first case every control register of a device has a number of input and output port. A device driver can read from a port and write to it with two `in` and `out` instructions which we already saw. If you want to know about currently registered port regions, you can learn about them by accessing `/proc/ioports`:

```
$ cat /proc/ioports
0000-0cf7 : PCI Bus 0000:00
  0000-001f : dma1
  0020-0021 : pic1
  0040-0043 : timer0
  0050-0053 : timer1
  0060-0060 : keyboard
  0064-0064 : keyboard
  0070-0077 : rtc0
  0080-008f : dma page reg
  00a0-00a1 : pic2
  00c0-00df : dma2
  00f0-00ff : fpu
    00f0-00f0 : PNP0C04:00
  03c0-03df : vesafb
  03f8-03ff : serial
  04d0-04d1 : pnp 00:06
  0800-087f : pnp 00:01
  0a00-0a0f : pnp 00:04
  0a20-0a2f : pnp 00:04
  0a30-0a3f : pnp 00:04
0cf8-0cff : PCI conf1
0d00-ffff : PCI Bus 0000:00
...
...
...
```

`/proc/ioports` provides information about which driver uses which address of a `I/O` port region. All of these memory regions, for example `0000-0cf7`, were claimed with the `request_region` function from the include/linux/ioport.h. Actually `request_region` is a macro which is defined as:

```
#define request_region(start,n,name)   __request_region(&ioport_resource, (start), (n), (name), 0)
```

As we can see it takes three parameters:

- `start` - begin of region;
- `n` - length of region;
- `name` - name of requester.

`request_region` allocates an `I/O` port region. Very often the `check_region` function is called before the `request_region` to check that the given address range is available and the `release_region` function to release the memory region. `request_region` returns a pointer to the `resource` structure. The `resource` structure represents an abstraction for a tree-like subset of system resources. We already saw the `resource` structure in the fifth part of the kernel initialization process and it looks as follows:

```
struct resource {
        resource_size_t start;
        resource_size_t end;
        const char *name;
        unsigned long flags;
        struct resource *parent, *sibling, *child;
};
```

and contains start and end addresses of the resource, the name, etc. Every `resource` structure contains pointers to the `parent` , `sibling` and `child` resources. As it has a parent and a child, it means that every subset of resources has root `resource` structure. For example, for `I/O` ports it is the `ioport_resource` structure:

```
struct resource ioport_resource = {
         .name   = "PCI IO",
         .start  = 0,
         .end    = IO_SPACE_LIMIT,
        .flags  = IORESOURCE_IO,
};
EXPORT_SYMBOL(ioport_resource);
```

Or for `iomem` , it is the `iomem_resource` structure:

```
struct resource iomem_resource = {
        .name   = "PCI mem",
        .start  = 0,
        .end    = -1,
        .flags  = IORESOURCE_MEM,
};
```

As I have mentioned before, `request_regions` is used to register I/O port regions and this macro is used in many places in the kernel. For example let's look at drivers/char/rtc.c. This source code file provides the Real Time Clock interface in the linux kernel. As every kernel module, `rtc` module contains `module_init` definition:

```
module_init(rtc_init);
```

where `rtc_init` is the `rtc` initialization function. This function is defined in the same `rtc.c` source code file. In the `rtc_init` function we can see a couple of calls to the `rtc_request_region` functions, which wrap `request_region` for example:

```
r = rtc_request_region(RTC_IO_EXTENT);
```

where `rtc_request_region` calls:

```
r = request_region(RTC_PORT(0), size, "rtc");
```

Here `RTC_IO_EXTENT` is the size of the memory region and it is `0x8` , `"rtc"` is the name of the region and `RTC_PORT` is:

```
#define RTC_PORT(x)     (0x70 + (x))
```

So with the `request_region(RTC_PORT(0), size, "rtc")` we register a memory region that starts at `0x70` and and has a size of `0x8` . Let's look at `/proc/ioports` :

```
~$ sudo cat /proc/ioports | grep rtc
0070-0077 : rtc0
```

So, we got it! Ok, that was it for the I/O ports. The second way to communicate with drivers is through the use of `I/O` memory. As I have mentioned above this works by mapping the control registers and the memory of a device to the memory address space. `I/O` memory is a set of contiguous addresses which are provided by a device to the CPU through a bus. None of the memory-mapped I/O addresses are used by the kernel directly. There is a special `ioremap` function which allows us to convert the physical address on a bus to a kernel virtual address. In other words, `ioremap` maps I/O physical memory regions to make them accessible from the kernel. The `ioremap` function takes two parameters:

- start of the memory region;
- size of the memory region;

The I/O memory mapping API provides functions to check, request and release memory regions as I/O memory. There are three functions for that:

- `request_mem_region`
- `release_mem_region`
- `check_mem_region`

```
~$ sudo cat /proc/iomem
...
...
...
be826000-be82cfff : ACPI Non-volatile Storage
be82d000-bf744fff : System RAM
bf745000-bfff4fff : reserved
bfff5000-dc041fff : System RAM
dc042000-dc0d2fff : reserved
dc0d3000-dc138fff : System RAM
dc139000-dc27dfff : ACPI Non-volatile Storage
dc27e000-deffefff : reserved
defff000-deffffff : System RAM
df000000-dfffffff : RAM buffer
e0000000-feafffff : PCI Bus 0000:00
  e0000000-efffffff : PCI Bus 0000:01
    e0000000-efffffff : 0000:01:00.0
  f7c00000-f7cfffff : PCI Bus 0000:06
    f7c00000-f7c0ffff : 0000:06:00.0
    f7c10000-f7c101ff : 0000:06:00.0
      f7c10000-f7c101ff : ahci
  f7d00000-f7dfffff : PCI Bus 0000:03
    f7d00000-f7d3ffff : 0000:03:00.0
      f7d00000-f7d3ffff : alx
...
...
...
```

Part of these addresses are from the call of the `e820_reserve_resources` function. We can find a call to this function in the arch/x86/kernel/setup.c and the function itself is defined in arch/x86/kernel/e820.c. `e820_reserve_resources` goes through the e820 map and inserts memory regions into the root `iomem` resource structure. All `e820` memory regions which are inserted into the `iomem` resource have the following types:

```
static inline const char *e820_type_to_string(int e820_type)
{
    switch (e820_type) {
    case E820_RESERVED_KERN:
```

```
        case E820_RAM:     return "System RAM";
        case E820_ACPI:     return "ACPI Tables";
        case E820_NVS:     return "ACPI Non-volatile Storage";
        case E820_UNUSABLE:     return "Unusable memory";
        default:     return "reserved";
        }
}
```

and we can see them in the `/proc/iomem` (read above).

Now let's try to understand how `ioremap` works. We already know a little about `ioremap`, we saw it in the fifth part about linux kernel initialization. If you have read this part, you can remember the call of the `early_ioremap_init` function from the arch/x86/mm/ioremap.c. Initialization of the `ioremap` is split into two parts: there is the early part which we can use before the normal `ioremap` is available and the normal `ioremap` which is available after `vmalloc` initialization and the call of `paging_init`. We do not know anything about `vmalloc` for now, so let's consider early initialization of the `ioremap`. First of all `early_ioremap_init` checks that `fixmap` is aligned on page middle directory boundary:

```
BUILD_BUG_ON((fix_to_virt(0) + PAGE_SIZE) & ((1 << PMD_SHIFT) - 1));
```

more about `BUILD_BUG_ON` you can read in the first part about Linux Kernel initialization. So `BUILD_BUG_ON` macro raises a compilation error if the given expression is true. In the next step after this check, we can see call of the `early_ioremap_setup` function from the mm/early_ioremap.c. This function presents generic initialization of the `ioremap`. `early_ioremap_setup` function fills the `slot_virt` array with the virtual addresses of the early fixmaps. All early fixmaps are after `__end_of_permanent_fixed_addresses` in memory. They start at `FIX_BITMAP_BEGIN` (top) and end with `FIX_BITMAP_END` (down). Actually there are `512` temporary boot-time mappings, used by early `ioremap`:

```
#define NR_FIX_BTMAPS        64
#define FIX_BTMAPS_SLOTS    8
#define TOTAL_FIX_BTMAPS    (NR_FIX_BTMAPS * FIX_BTMAPS_SLOTS)
```

and `early_ioremap_setup`:

```
void __init early_ioremap_setup(void)
{
        int i;

        for (i = 0; i < FIX_BTMAPS_SLOTS; i++)
                if (WARN_ON(prev_map[i]))
                        break;

        for (i = 0; i < FIX_BTMAPS_SLOTS; i++)
                slot_virt[i] = __fix_to_virt(FIX_BTMAP_BEGIN - NR_FIX_BTMAPS*i);
}
```

the `slot_virt` and other arrays are defined in the same source code file:

```
static void __iomem *prev_map[FIX_BTMAPS_SLOTS] __initdata;
static unsigned long prev_size[FIX_BTMAPS_SLOTS] __initdata;
static unsigned long slot_virt[FIX_BTMAPS_SLOTS] __initdata;
```

`slot_virt` contains the virtual addresses of the `fix-mapped` areas, `prev_map` array contains addresses of the early ioremap areas. Note that I wrote above: `Actually there are 512 temporary boot-time mappings, used by early ioremap` and you can see that all arrays are defined with the `__initdata` attribute which means that this memory will be released after the kernel initialization process. After `early_ioremap_setup` has finished its work, we're getting page middle directory where early ioremap begins with the `early_ioremap_pmd` function which just gets the base address of the page global directory and calculates the page middle directory for the given address:

```
static inline pmd_t * __init early_ioremap_pmd(unsigned long addr)
{
    pgd_t *base = __va(read_cr3());
    pgd_t *pgd = &base[pgd_index(addr)];
    pud_t *pud = pud_offset(pgd, addr);
    pmd_t *pmd = pmd_offset(pud, addr);
    return pmd;
}
```

After this we fill `bm_pte` (early ioremap page table entries) with zeros and call the `pmd_populate_kernel` function:

```
pmd = early_ioremap_pmd(fix_to_virt(FIX_BTMAP_BEGIN));
memset(bm_pte, 0, sizeof(bm_pte));
pmd_populate_kernel(&init_mm, pmd, bm_pte);
```

`pmd_populate_kernel` takes three parameters:

- `init_mm` - memory descriptor of the `init` process (you can read about it in the previous part);
- `pmd` - page middle directory of the beginning of the `ioremap` fixmaps;
- `bm_pte` - early `ioremap` page table entries array which defined as:

```
static pte_t bm_pte[PAGE_SIZE/sizeof(pte_t)] __page_aligned_bss;
```

The `pmd_populate_kernel` function is defined in the arch/x86/include/asm/pgalloc.h and populates the page middle directory (`pmd`) provided as an argument with the given page table entries (`bm_pte`):

```
static inline void pmd_populate_kernel(struct mm_struct *mm,
                                       pmd_t *pmd, pte_t *pte)
{
        paravirt_alloc_pte(mm, __pa(pte) >> PAGE_SHIFT);
        set_pmd(pmd, __pmd(__pa(pte) | _PAGE_TABLE));
}
```

where `set_pmd` is:

```
#define set_pmd(pmdp, pmd)              native_set_pmd(pmdp, pmd)
```

and `native_set_pmd` is:

```
static inline void native_set_pmd(pmd_t *pmdp, pmd_t pmd)
{
        *pmdp = pmd;
}
```

That's all. Early `ioremap` is ready to use. There are a couple of checks in the `early_ioremap_init` function, but they are not so important, anyway initialization of the `ioremap` is finished.

Use of early ioremap

As soon as early `ioremap` has been setup successfully, we can use it. It provides two functions:

- early_ioremap
- early_iounmap

for mapping/unmapping of I/O physical address to virtual address. Both functions depend on the `CONFIG_MMU` configuration option. Memory management unit is a special block of memory management. The main purpose of this block is the translation of physical addresses to virtual addresses. The memory management unit knows about the high-level page table addresses (`pgd`) from the `cr3` control register. If `CONFIG_MMU` options is set to `n` , `early_ioremap` just returns the given physical address and `early_iounmap` does nothing. If `CONFIG_MMU` option is set to `y` , `early_ioremap` calls `__early_ioremap` which takes three parameters:

- `phys_addr` - base physical address of the `I/O` memory region to map on virtual addresses;
- `size` - size of the `I/O` memory region;
- `prot` - page table entry bits.

First of all in the `__early_ioremap` , we go through all early ioremap fixmap slots and search for the first free one in the `prev_map` array. When we found it we remember its number in the `slot` variable and set up size:

```
slot = -1;
for (i = 0; i < FIX_BTMAPS_SLOTS; i++) {
    if (!prev_map[i]) {
        slot = i;
        break;
    }
}
...
...
...
prev_size[slot] = size;
last_addr = phys_addr + size - 1;
```

In the next spte we can see the following code:

```
offset = phys_addr & ~PAGE_MASK;
phys_addr &= PAGE_MASK;
size = PAGE_ALIGN(last_addr + 1) - phys_addr;
```

Here we are using `PAGE_MASK` for clearing all bits in the `phys_addr` except the first 12 bits. `PAGE_MASK` macro is defined as:

```
#define PAGE_MASK       (~(PAGE_SIZE-1))
```

We know that size of a page is 4096 bytes or `1000000000000` in binary. `PAGE_SIZE - 1` will be `111111111111` , but with `~` , we will get `000000000000` , but as we use `~PAGE_MASK` we will get `111111111111` again. On the second line we do the same but clear the first 12 bits and getting page-aligned size of the area on the third line. We getting aligned area and now we need to get the number of pages which are occupied by the new `ioremap` area and calculate the fix-mapped index from `fixed_addresses` in the next steps:

```
nrpages = size >> PAGE_SHIFT;
idx = FIX_BTMAP_BEGIN - NR_FIX_BTMAPS*slot;
```

Now we can fill `fix-mapped` area with the given physical addresses. On every iteration in the loop, we call the `__early_set_fixmap` function from the arch/x86/mm/ioremap.c, increase the given physical address by the page size which is `4096` bytes and update the `addresses` index and the number of pages:

```
while (nrpages > 0) {
    __early_set_fixmap(idx, phys_addr, prot);
    phys_addr += PAGE_SIZE;
    --idx;
    --nrpages;
}
```

The `__early_set_fixmap` function gets the page table entry (stored in the `bm_pte`, see above) for the given physical address with:

```
pte = early_ioremap_pte(addr);
```

In the next step of `early_ioremap_pte` we check the given page flags with the `pgprot_val` macro and call `set_pte` or `pte_clear` depending on the flags given:

```
if (pgprot_val(flags))
        set_pte(pte, pfn_pte(phys >> PAGE_SHIFT, flags));
    else
        pte_clear(&init_mm, addr, pte);
```

As you can see above, we passed `FIXMAP_PAGE_IO` as flags to the `__early_ioremap`. `FIXMPA_PAGE_IO` expands to the:

```
(__PAGE_KERNEL_EXEC | _PAGE_NX)
```

flags, so we call `set_pte` function to set the page table entry which works in the same manner as `set_pmd` but for PTEs (read above about it). As we have set all `PTEs` in the loop, we can now take a look at the call of the `__flush_tlb_one` function:

```
__flush_tlb_one(addr);
```

This function is defined in arch/x86/include/asm/tlbflush.h and calls `__flush_tlb_single` or `__flush_tlb` depending on the value of `cpu_has_invlpg`:

```
static inline void __flush_tlb_one(unsigned long addr)
{
        if (cpu_has_invlpg)
                __flush_tlb_single(addr);
        else
                __flush_tlb();
}
```

The `__flush_tlb_one` function invalidates the given address in the TLB. As you just saw we updated the paging structure, but `TLB` is not informed of the changes, that's why we need to do it manually. There are two ways to do it. The first is to update the `cr3` control register and the `__flush_tlb` function does this:

```
native_write_cr3(native_read_cr3());
```

The second method is to use the `invlpg` instruction to invalidate the `TLB` entry. Let's look at the `__flush_tlb_one` implementation. As you can see, first of all the function checks `cpu_has_invlpg` which is defined as:

```
#if defined(CONFIG_X86_INVLPG) || defined(CONFIG_X86_64)
# define cpu_has_invlpg         1
#else
# define cpu_has_invlpg         (boot_cpu_data.x86 > 3)
#endif
```

If a CPU supports the `invlpg` instruction, we call the `__flush_tlb_single` macro which expands to the call of `__native_flush_tlb_single`:

```
static inline void __native_flush_tlb_single(unsigned long addr)
{
        asm volatile("invlpg (%0)" ::"r" (addr) : "memory");
}
```

or call `__flush_tlb` which just updates the `cr3` register as we have seen. After this step execution of the `__early_set_fixmap` function is finished and we can go back to the `__early_ioremap` implementation. When we have set up the fixmap area for the given address, we need to save the base virtual address of the I/O Re-mapped area in the `prev_map` using the `slot` index:

```
prev_map[slot] = (void __iomem *)(offset + slot_virt[slot]);
```

and return it.

The second function, `early_iounmap`, unmaps an `I/O` memory region. This function takes two parameters: base address and size of a `I/O` region and generally looks very similar to `early_ioremap`. It also goes through fixmap slots and looks for a slot with the given address. After that, it gets the index of the fixmap slot and calls `__late_clear_fixmap` or `__early_set_fixmap` depending on the `after_paging_init` value. It calls `__early_set_fixmap` with one difference to how `early_ioremap` does it: `early_iounmap` passes `zero` as physical address. And in the end it sets the address of the I/O memory region to `NULL`:

```
prev_map[slot] = NULL;
```

That's all about `fixmaps` and `ioremap`. Of course this part does not cover all features of `ioremap`, only early ioremap but there is also normal ioremap. But we need to know more things before we study that in more detail.

So, this is the end!

Conclusion

This is the end of the second part about linux kernel memory management. If you have questions or suggestions, ping me on twitter 0xAX, drop me an email or just create an issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me a PR to linux-insides.

Links

- apic
- vsyscall
- Intel Trusted Execution Technology
- Xen
- Real Time Clock
- e820
- Memory management unit
- TLB
- Paging
- Linux kernel memory management Part 1.

Linux kernel memory management Part 3.

Introduction to the kmemcheck in the Linux kernel

This is the third part of the chapter which describes memory management in the Linux kernel and in the previous part of this chapter we met two memory management related concepts:

- `Fix-Mapped Addresses` ;
- `ioremap` .

The first concept represents special area in virtual memory, whose corresponding physical mapping is calculated in compile-time. The second concept provides ability to map input/output related memory to virtual memory.

For example if you will look at the output of the `/proc/iomem` :

```
$ sudo cat /proc/iomem

00000000-00000fff : reserved
00001000-0009d7ff : System RAM
0009d800-0009ffff : reserved
000a0000-000bffff : PCI Bus 0000:00
000c0000-000cffff : Video ROM
000d0000-000d3fff : PCI Bus 0000:00
000d4000-000d7fff : PCI Bus 0000:00
000d8000-000dbfff : PCI Bus 0000:00
000dc000-000dffff : PCI Bus 0000:00
000e0000-000fffff : reserved
...
...
...
```

you will see map of the system's memory for each physical device. Here the first column displays the memory registers used by each of the different types of memory. The second column lists the kind of memory located within those registers. Or for example:

```
$ sudo cat /proc/ioports

0000-0cf7 : PCI Bus 0000:00
  0000-001f : dma1
  0020-0021 : pic1
  0040-0043 : timer0
  0050-0053 : timer1
  0060-0060 : keyboard
  0064-0064 : keyboard
  0070-0077 : rtc0
  0080-008f : dma page reg
  00a0-00a1 : pic2
  00c0-00df : dma2
  00f0-00ff : fpu
    00f0-00f0 : PNP0C04:00
  03c0-03df : vga+
  03f8-03ff : serial
  04d0-04d1 : pnp 00:06
  0800-087f : pnp 00:01
  0a00-0a0f : pnp 00:04
  0a20-0a2f : pnp 00:04
  0a30-0a3f : pnp 00:04
...
...
...
```

can show us lists of currently registered port regions used for input or output communication with a device. All memory-mapped I/O addresses are not used by the kernel directly. So, before the Linux kernel can use such memory, it must map it to the virtual memory space which is the main purpose of the `ioremap` mechanism. Note that we saw only early `ioremap` in the previous part. Soon we will look at the implementation of the non-early `ioremap` function. But before this we must learn other things, like a different types of memory allocators and etc., because in other way it will be very difficult to understand it.

So, before we will move on to the non-early memory management of the Linux kernel, we will see some mechanisms which provide special abilities for debugging, check of memory leaks, memory control and etc. It will be easier to understand how memory management arranged in the Linux kernel after learning of all of these things.

As you already may guess from the title of this part, we will start to consider memory mechanisms from the kmemcheck. As we always did in other chapters, we will start to consider from theoretical side and will learn what is `kmemcheck` mechanism in general and only after this, we will see how it is implemented in the Linux kernel.

So let's start. What is it `kmemcheck` in the Linux kernel? As you may guess from the name of this mechanism, the `kmemcheck` checks memory. That's true. Main point of the `kmemcheck` mechanism is to check that some kernel code accesses `uninitialized memory`. Let's take following simple C program:

```
#include <stdlib.h>
#include <stdio.h>

struct A {
        int a;
};

int main(int argc, char **argv) {
        struct A *a = malloc(sizeof(struct A));
        printf("a->a = %d\n", a->a);
        return 0;
}
```

Here we allocate memory for the `A` structure and tries to print value of the `a` field. If we will compile this program without additional options:

```
gcc test.c -o test
```

The compiler will not show us warning that `a` filed is not unitialized. But if we will run this program with valgrind tool, we will see the following output:

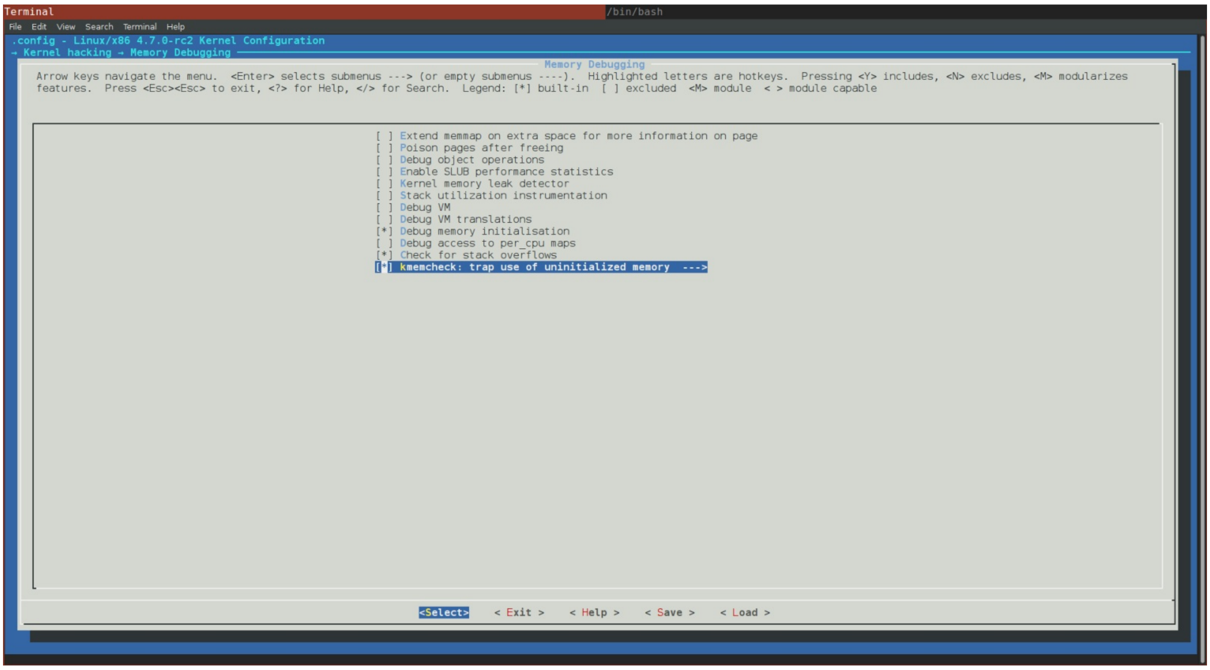
```
~$   valgrind --leak-check=yes ./test
==28469== Memcheck, a memory error detector
==28469== Copyright (C) 2002-2015, and GNU GPL'd, by Julian Seward et al.
==28469== Using Valgrind-3.11.0 and LibVEX; rerun with -h for copyright info
==28469== Command: ./test
==28469==
==28469== Conditional jump or move depends on uninitialised value(s)
==28469==    at 0x4E820EA: vfprintf (in /usr/lib64/libc-2.22.so)
==28469==    by 0x4E88D48: printf (in /usr/lib64/libc-2.22.so)
==28469==    by 0x4005B9: main (in /home/alex/test)
==28469==
==28469== Use of uninitialised value of size 8
==28469==    at 0x4E7E0BB: _itoa_word (in /usr/lib64/libc-2.22.so)
==28469==    by 0x4E8262F: vfprintf (in /usr/lib64/libc-2.22.so)
==28469==    by 0x4E88D48: printf (in /usr/lib64/libc-2.22.so)
==28469==    by 0x4005B9: main (in /home/alex/test)
...
...
...
```

Actually the `kmemcheck` mechanism does the same for the kernel, what the `valgrind` does for userspace programs. It check uninitialized memory.

To enable this mechanism in the Linux kernel, you need to enable the `CONFIG_KMEMCHECK` kernel configuration option in the:

```
Kernel hacking
  -> Memory Debugging
```

menu of the Linux kernel configuration:



We may not only enable support of the `kmemcheck` mechanism in the Linux kernel, but it also provides some configuration options for us. We will see all of these options in the next paragraph of this part. Last note before we will consider how does the `kmemcheck` check memory. Now this mechanism is implemented only for the x86_64 architecture. You can be sure if you will look in the arch/x86/Kconfig `x86` related kernel configuration file, you will see following lines:

```
config X86
  ...
  ...
  ...
  select HAVE_ARCH_KMEMCHECK
  ...
  ...
  ...
```

So, there is no anything which is specific for other architectures.

Ok, so we know that `kmemcheck` provides mechanism to check usage of `uninitialized memory` in the Linux kernel and how to enable it. How it does these checks? When the Linux kernel tries to allocate some memory i.e. something is called like this:

```
struct my_struct *my_struct = kmalloc(sizeof(struct my_struct), GFP_KERNEL);
```

or in other words somebody wants to access a page, a page fault exception is generated. This is achieved by the fact that the `kmemcheck` marks memory pages as `non-present` (more about this you can read in the special part which is devoted to Paging). If a `page fault` exception is occurred, the exception handler knows about it and in a case when the `kmemcheck` is enabled it transfers control to it. After the `kmemcheck` will finish its checks, the page will be marked as `present` and the interrupted code will be able to continue execution. There is little subtlety in this chain. When the first instruction of interrupted code will be executed, the `kmemcheck` will mark the page as `non-present` again. In this way next access to memory will be caught again.

We just considered the `kmemcheck` mechanism from theoretical side. Now let's consider how it is implemented in the Linux kernel.

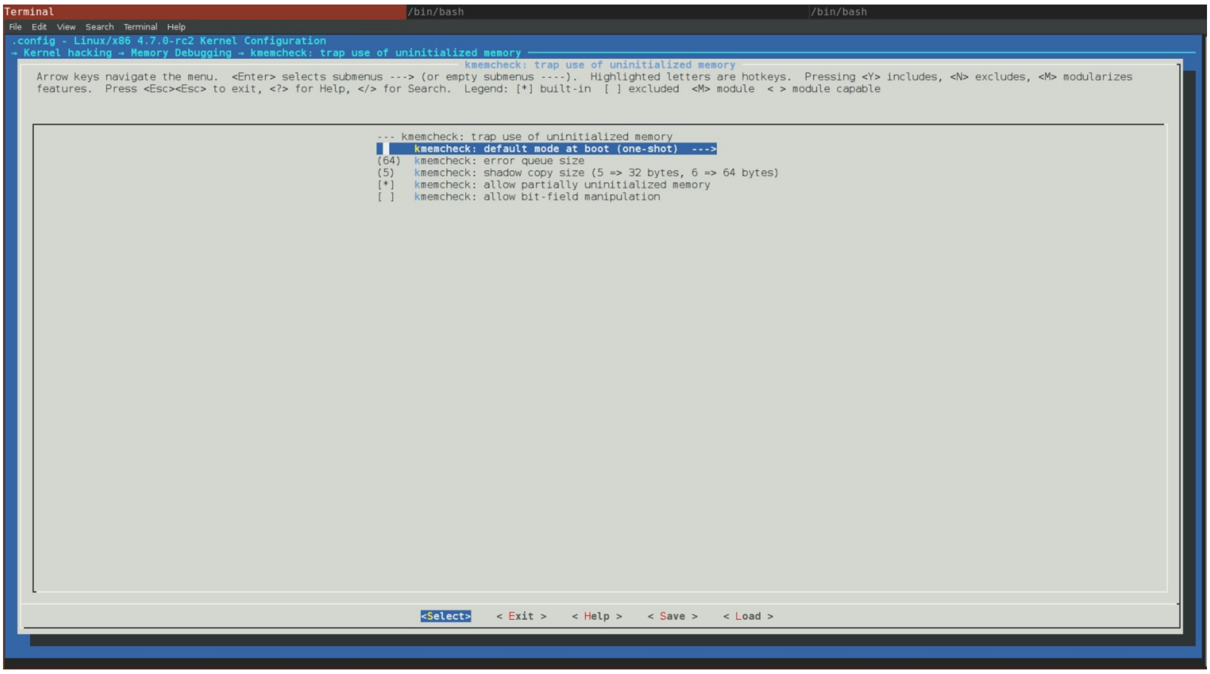
Implementation of the `kmemcheck` mechanism in the Linux kernel

So, now we know what is it `kmemcheck` and what it does in the Linux kernel. Time to see at its implementation in the Linux kernel. Implementation of the `kmemcheck` is split in two parts. The first is generic part is located in the mm/kmemcheck.c source code file and the second x86_64 architecture-specific part is located in the arch/x86/mm/kmemcheck directory.

Let's start from the initialization of this mechanism. We already know that to enable the `kmemcheck` mechanism in the Linux kernel, we must enable the `CONFIG_KMEMCHECK` kernel configuration option. But besides this, we need to pass one of following parameters:

- kmemcheck=0 (disabled)
- kmemcheck=1 (enabled)
- kmemcheck=2 (one-shot mode)

to the Linux kernel command line. The first two are clear, but the last needs a little explanation. This option switches the `kmemcheck` in a special mode when it will be turned off after detecting the first use of uninitialized memory. Actually this mode is enabled by default in the Linux kernel:



We know from the seventh part of the chapter which describes initialization of the Linux kernel that the kernel command line is parsed during initialization of the Linux kernel in `do_initcall_level`, `do_early_param` functions. Actually the `kmemcheck` subsystem consists from two stages. The first stage is early. If we will look at the mm/kmemcheck.c source code file, we will see the `param_kmemcheck` function which is will be called during early command line parsing:

```
static int __init param_kmemcheck(char *str)
{
    int val;
    int ret;
```

```
    if (!str)
        return -EINVAL;

    ret = kstrtoint(str, 0, &val);
    if (ret)
        return ret;
    kmemcheck_enabled = val;
    return 0;
}

early_param("kmemcheck", param_kmemcheck);
```

As we already saw, the `param_kmemcheck` may have one of the following values: `0` (enabled), `1` (disabled) or `2` (one-shot). The implementation of the `param_kmemcheck` is pretty simple. We just convert string value of the `kmemcheck` command line option to integer representation and set it to the `kmemcheck_enabled` variable.

The second stage will be executed during initialization of the Linux kernel, rather during initialization of early initcalls. The second stage is represented by the `kmemcheck_init` :

```
int __init kmemcheck_init(void)
{
    ...
    ...
    ...
}

early_initcall(kmemcheck_init);
```

Main goal of the `kmemcheck_init` function is to call the `kmemcheck_selftest` function and check its result:

```
if (!kmemcheck_selftest()) {
    printk(KERN_INFO "kmemcheck: self-tests failed; disabling\n");
    kmemcheck_enabled = 0;
    return -EINVAL;
}

printk(KERN_INFO "kmemcheck: Initialized\n");
```

and return with the `EINVAL` if this check is failed. The `kmemcheck_selftest` function checks sizes of different memory access related opcodes like `rep movsb` , `movzwq` and etc. If sizes of opcodes are equal to expected sizes, the `kmemcheck_selftest` will return `true` and `false` in other way.

So when the somebody will call:

```
struct my_struct *my_struct = kmalloc(sizeof(struct my_struct), GFP_KERNEL);
```

through a series of different function calls the `kmem_getpages` function will be called. This function is defined in the mm/slab.c source code file and main goal of this function tries to allocate pages with the given flags. In the end of this function we can see following code:

```
if (kmemcheck_enabled && !(cachep->flags & SLAB_NOTRACK)) {
    kmemcheck_alloc_shadow(page, cachep->gfporder, flags, nodeid);

    if (cachep->ctor)
        kmemcheck_mark_uninitialized_pages(page, nr_pages);
    else
        kmemcheck_mark_unallocated_pages(page, nr_pages);
}
```

So, here we check that the if `kmemcheck` is enabled and the `SLAB_NOTRACK` bit is not set in flags we set `non-present` bit for the just allocated page. The `SLAB_NOTRACK` bit tell us to not track uninitialized memory. Additionally we check if a cache object has constructor (details will be considered in next parts) we mark allocated page as uninitialized or unallocated in other way. The `kmemcheck_alloc_shadow` function is defined in the mm/kmemcheck.c source code file and does following things:

```
void kmemcheck_alloc_shadow(struct page *page, int order, gfp_t flags, int node)
{
    struct page *shadow;

       shadow = alloc_pages_node(node, flags | __GFP_NOTRACK, order);

       for(i = 0; i < pages; ++i)
        page[i].shadow = page_address(&shadow[i]);

       kmemcheck_hide_pages(page, pages);
}
```

First of all it allocates memory space for the shadow bits. If this bit is set in a page, this means that this page is tracked by the `kmemcheck`. After we allocated space for the shadow bit, we fill all allocated pages with this bit. In the end we just call the `kmemcheck_hide_pages` function with the pointer to the allocated page and number of these pages. The `kmemcheck_hide_pages` is architecture-specific function, so its implementation is located in the arch/x86/mm/kmemcheck/kmemcheck.c source code file. The main goal of this function is to set `non-present` bit in given pages. Let's look at the implementation of this function:

```
void kmemcheck_hide_pages(struct page *p, unsigned int n)
{
    unsigned int i;

    for (i = 0; i < n; ++i) {
        unsigned long address;
        pte_t *pte;
        unsigned int level;

        address = (unsigned long) page_address(&p[i]);
        pte = lookup_address(address, &level);
        BUG_ON(!pte);
        BUG_ON(level != PG_LEVEL_4K);

        set_pte(pte, __pte(pte_val(*pte) & ~_PAGE_PRESENT));
        set_pte(pte, __pte(pte_val(*pte) | _PAGE_HIDDEN));
        __flush_tlb_one(address);
    }
}
```

Here we go through all pages and and tries to get `page table entry` for each page. If this operation was successful, we unset present bit and set hidden bit in each page. In the end we flush translation lookaside buffer, because some pages was changed. From this point allocated pages are tracked by the `kmemcheck`. Now, as `present` bit is unset, the page fault execution will be occurred right after the `kmalloc` will return pointer to allocated space and a code will try to access this memory.

As you may remember from the second part of the Linux kernel initialization chapter, the `page fault` handler is located in the arch/x86/mm/fault.c source code file and represented by the `do_page_fault` function. We can see following check from the beginning of the `do_page_fault` function:

```
static noinline void
__do_page_fault(struct pt_regs *regs, unsigned long error_code,
        unsigned long address)
{
    ...
    ...
    ...
    if (kmemcheck_active(regs))
        kmemcheck_hide(regs);
```

```
    ...
    ...
    ...
}
```

The `kmemcheck_active` gets `kmemcheck_context` per-cpu structure and return the result of comparison of the `balance` field of this structure with zero:

```
bool kmemcheck_active(struct pt_regs *regs)
{
    struct kmemcheck_context *data = this_cpu_ptr(&kmemcheck_context);

    return data->balance > 0;
}
```

The `kmemcheck_context` is structure which describes current state of the `kmemcheck` mechanism. It stored uninitialized addresses, number of such addresses and etc. The `balance` field of this structure represents current state of the `kmemcheck` or in other words it can tell us did `kmemcheck` already hid pages or not yet. If the `data->balance` is greater than zero, the `kmemcheck_hide` function will be called. This means than `kmemecheck` already set `present` bit for given pages and now we need to hide pages again to cause next step to page fault. This function will hide addresses of pages again by unsetting of `present` bit. This means that one session of `kmemcheck` already finished and new page fault occurred. At the first step the `kmemcheck_active` will return false as the `data->balance` is zero for the start and the `kmemcheck_hide` will not be called. Next, we may see following line of code in the `do_page_fault` :

```
if (kmemcheck_fault(regs, address, error_code))
        return;
```

First of all the `kmemcheck_fault` function checks that the fault was occured by the correct reason. At first we check the flags register and check that we are in normal kernel mode:

```
if (regs->flags & X86_VM_MASK)
        return false;
if (regs->cs != __KERNEL_CS)
        return false;
```

If these checks wasn't successful we return from the `kmemcheck_fault` function as it was not `kmemcheck` related page fault. After this we try to lookup a `page table entry` related to the faulted address and if we can't find it we return:

```
pte = kmemcheck_pte_lookup(address);
if (!pte)
    return false;
```

Last two steps of the `kmemcheck_fault` function is to call the `kmemcheck_access` function which check access to the given page and show addresses again by setting present bit in the given page. The `kmemcheck_access` function does all main job. It check current instruction which caused a page fault. If it will find an error, the context of this error will be saved by `kmemcheck` to the ring queue:

```
static struct kmemcheck_error error_fifo[CONFIG_KMEMCHECK_QUEUE_SIZE];
```

The `kmemcheck` mechanism declares special tasklet:

```
static DECLARE_TASKLET(kmemcheck_tasklet, &do_wakeup, 0);
```

which runs the `do_wakeup` function from the arch/x86/mm/kmemcheck/error.c source code file when it will be scheduled to run.

The `do_wakeup` function will call the `kmemcheck_error_recall` function which will print errors collected by `kmemcheck`. As we already saw the:

```
kmemcheck_show(regs);
```

function will be called in the end of the `kmemcheck_fault` function. This function will set present bit for the given pages again:

```
if (unlikely(data->balance != 0)) {
    kmemcheck_show_all();
    kmemcheck_error_save_bug(regs);
    data->balance = 0;
    return;
}
```

Where the `kmemcheck_show_all` function calls the `kmemcheck_show_addr` for each address:

```
static unsigned int kmemcheck_show_all(void)
{
    struct kmemcheck_context *data = this_cpu_ptr(&kmemcheck_context);
    unsigned int i;
    unsigned int n;

    n = 0;
    for (i = 0; i < data->n_addrs; ++i)
        n += kmemcheck_show_addr(data->addr[i]);

    return n;
}
```

by the call of the `kmemcheck_show_addr` :

```
int kmemcheck_show_addr(unsigned long address)
{
    pte_t *pte;

    pte = kmemcheck_pte_lookup(address);
    if (!pte)
        return 0;

    set_pte(pte, __pte(pte_val(*pte) | _PAGE_PRESENT));
    __flush_tlb_one(address);
    return 1;
}
```

In the end of the `kmemcheck_show` function we set the TF flag if it wasn't set:

```
if (!(regs->flags & X86_EFLAGS_TF))
    data->flags = regs->flags;
```

We need to do it because we need to hide pages again after first executed instruction after a page fault will be handled. In a case when the `TF` flag, so the processor will switch into single-step mode after the first instruction will be executed. In this case `debug` exception will occurred. From this moment pages will be hidden again and execution will be continued. As pages hidden from this moment, page fault exception will occur again and `kmemcheck` continue to check/collect errors again and print them from time to time.

That's all.

Conclusion

This is the end of the third part about linux kernel memory management. If you have questions or suggestions, ping me on twitter 0xAX, drop me an email or just create an issue. In the next part we will see yet another memory debugging related tool - `kmemleak`.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me a PR to linux-insides.

Links

- memory management
- debugging
- memory leaks
- kmemcheck documentation
- valgrind
- Paging
- page fault
- initcalls
- opcode
- translation lookaside buffer
- per-cpu variables
- flags register
- tasklet
- Previous part

Cgroups

This chapter describes `control groups` mechanism in the Linux kernel.

Control Groups

Introduction

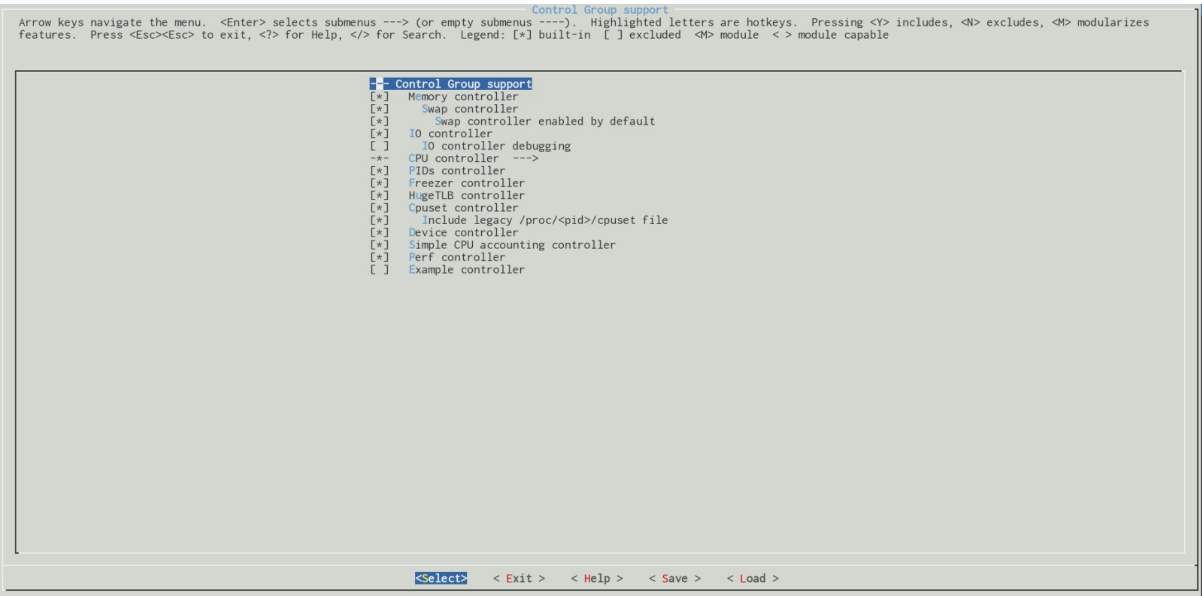
This is the first part of the new chapter of the linux insides book and as you may guess by part's name - this part will cover control groups or `cgroups` mechanism in the Linux kernel.

`Cgroups` are special mechanism provided by the Linux kernel which allows us to allocate kind of `resources` like processor time, number of processes per group, amount of memory per control group or combination of such resources for a process or set of processes. `Cgroups` are organized hierarchically and here this mechanism is similar to usual processes as they are hierarchical too and child `cgroups` inherit set of certain parameters from their parents. But actually they are not the same. The main differences between `cgroups` and normal processes that many different hierarchies of control groups may exist simultaneously in one time while normal process tree is always single. This was not a casual step because each control group hierarchy is attached to set of control group `subsystems` .

One `control group subsystem` represents one kind of resources like a processor time or number of pids or in other words number of processes for a `control group` . Linux kernel provides support for following twelve `control group subsystems` :

- `cpuset` - assigns individual processor(s) and memory nodes to task(s) in a group;
- `cpu` - uses the scheduler to provide cgroup tasks access to the processor resources;
- `cpuacct` - generates reports about processor usage by a group;
- `io` - sets limit to read/write from/to block devices;
- `memory` - sets limit on memory usage by a task(s) from a group;
- `devices` - allows access to devices by a task(s) from a group;
- `freezer` - allows to suspend/resume for a task(s) from a group;
- `net_cls` - allows to mark network packets from task(s) from a group;
- `net_prio` - provides a way to dynamically set the priority of network traffic per network interface for a group;
- `perf_event` - provides access to perf events) to a group;
- `hugetlb` - activates support for huge pages for a group;
- `pid` - sets limit to number of processes in a group.

Each of these control group subsystems depends on related configuration option. For example the `cpuset` subsystem should be enabled via `CONFIG_CPUSETS` kernel configuration option, the `io` subsystem via `CONFIG_BLK_CGROUP` kernel configuration option and etc. All of these kernel configuration options may be found in the `General setup → Control Group support` menu:



You may see enabled control groups on your computer via proc filesystem:

```
$ cat /proc/cgroups
#subsys_name    hierarchy    num_cgroups    enabled
cpuset    8    1    1
cpu    7    66    1
cpuacct    7    66    1
blkio    11    66    1
memory    9    94    1
devices    6    66    1
freezer    2    1    1
net_cls    4    1    1
perf_event    3    1    1
net_prio    4    1    1
hugetlb    10    1    1
pids    5    69    1
```

or via sysfs:

```
$ ls -l /sys/fs/cgroup/
total 0
dr-xr-xr-x 5 root root  0 Dec  2 22:37 blkio
lrwxrwxrwx 1 root root 11 Dec  2 22:37 cpu -> cpu,cpuacct
lrwxrwxrwx 1 root root 11 Dec  2 22:37 cpuacct -> cpu,cpuacct
dr-xr-xr-x 5 root root  0 Dec  2 22:37 cpu,cpuacct
dr-xr-xr-x 2 root root  0 Dec  2 22:37 cpuset
dr-xr-xr-x 5 root root  0 Dec  2 22:37 devices
dr-xr-xr-x 2 root root  0 Dec  2 22:37 freezer
dr-xr-xr-x 2 root root  0 Dec  2 22:37 hugetlb
dr-xr-xr-x 5 root root  0 Dec  2 22:37 memory
lrwxrwxrwx 1 root root 16 Dec  2 22:37 net_cls -> net_cls,net_prio
dr-xr-xr-x 2 root root  0 Dec  2 22:37 net_cls,net_prio
lrwxrwxrwx 1 root root 16 Dec  2 22:37 net_prio -> net_cls,net_prio
dr-xr-xr-x 2 root root  0 Dec  2 22:37 perf_event
dr-xr-xr-x 5 root root  0 Dec  2 22:37 pids
dr-xr-xr-x 5 root root  0 Dec  2 22:37 systemd
```

As you already may guess that `control groups` mechanism is not such mechanism which was invented only directly to the needs of the Linux kernel, but mostly for userspace needs. To use a `control group`, we should create it at first. We may create a `cgroup` via two ways.

The first way is to create subdirectory in any subsystem from `sys/fs/cgroup` and add a pid of a task to a `tasks` file which will be created automatically right after we will create the subdirectory.

The second way is to create/destroy/manage `cgroups` with utils from `libcgroup` library (`libcgroup-tools` in Fedora).

Let's consider simple example. Following bash script will print a line to `/dev/tty` device which represents control terminal for the current process:

```
#!/bin/bash

while :
do
    echo "print line" > /dev/tty
    sleep 5
done
```

So, if we will run this script we will see following result:

```
$ sudo chmod +x cgroup_test_script.sh
~$ ./cgroup_test_script.sh
print line
print line
print line
...
...
...
```

Now let's go to the place where `cgroupfs` is mounted on our computer. As we just saw, this is `/sys/fs/cgroup` directory, but you may mount it everywhere you want.

```
$ cd /sys/fs/cgroup
```

And now let's go to the `devices` subdirectory which represents kind of resources that allows or denies access to devices by tasks in a `cgroup` :

```
# cd /devices
```

and create `cgroup_test_group` directory there:

```
# mkdir cgroup_test_group
```

After creation of the `cgroup_test_group` directory, following files will be generated there:

```
/sys/fs/cgroup/devices/cgroup_test_group$ ls -l
total 0
-rw-r--r-- 1 root root 0 Dec  3 22:55 cgroup.clone_children
-rw-r--r-- 1 root root 0 Dec  3 22:55 cgroup.procs
--w------- 1 root root 0 Dec  3 22:55 devices.allow
--w------- 1 root root 0 Dec  3 22:55 devices.deny
-r--r--r-- 1 root root 0 Dec  3 22:55 devices.list
-rw-r--r-- 1 root root 0 Dec  3 22:55 notify_on_release
-rw-r--r-- 1 root root 0 Dec  3 22:55 tasks
```

For this moment we are interested in `tasks` and `devices.deny` files. The first `tasks` files should contain pid(s) of processes which will be attached to the `cgroup_test_group` . The second `devices.deny` file contain list of denied devices. By default a newly created group has no any limits for devices access. To forbid a device (in our case it is `/dev/tty`) we should write to the `devices.deny` following line:

```
# echo "c 5:0 w" > devices.deny
```

Let's go step by step through this line. The first `c` letter represents type of a device. In our case the `/dev/tty` is `char device`. We can verify this from output of `ls` command:

```
~$ ls -l /dev/tty
crw-rw-rw- 1 root tty 5, 0 Dec  3 22:48 /dev/tty
```

see the first `c` letter in a permissions list. The second part is `5:0` is minor and major numbers of the device. You can see these numbers in the output of `ls` too. And the last `w` letter forbids tasks to write to the specified device. So let's start the `cgroup_test_script.sh` script:

```
~$ ./cgroup_test_script.sh
print line
print line
print line
...
...
```

and add pid of this process to the `devices/tasks` file of our group:

```
# echo $(pidof -x cgroup_test_script.sh) > /sys/fs/cgroup/devices/cgroup_test_group/tasks
```

The result of this action will be as expected:

```
~$ ./cgroup_test_script.sh
print line
print line
print line
print line
print line
print line
./cgroup_test_script.sh: line 5: /dev/tty: Operation not permitted
```

Similar situation will be when you will run you docker) containers for example:

```
~$ docker ps
CONTAINER ID        IMAGE               COMMAND                  CREATED             STATUS              PORTS
                   NAMES
fa2d2085cd1c        mariadb:10          "docker-entrypoint..."   12 days ago         Up 4 minutes        0.0.0.
0:3306->3306/tcp   mysql-work

~$ cat /sys/fs/cgroup/devices/docker/fa2d2085cd1c8d797002c77387d2061f56fefb470892f140d0dc511bd4d9bb61/tasks | h
ead -3
5501
5584
5585
...
...
...
```

So, during startup of a `docker` container, `docker` will create a `cgroup` for processes in this container:

```
$ docker exec -it mysql-work /bin/bash
$ top
  PID USER      PR  NI    VIRT    RES    SHR S  %CPU %MEM     TIME+ COMMAND
                                               1 mysql     20   0  963996 101268  15744 S   0.0  0.6   0:00.46
mysqld
```

```
    71 root      20   0   20248   3028   2732 S   0.0  0.0   0:00.01 bash
    77 root      20   0   21948   2424   2056 R   0.0  0.0   0:00.00 top
```

And we may see this `cgroup` on host machine:

```
$ systemd-cgls

Control group /:
-.slice
├─docker
│ └─fa2d2085cd1c8d797002c77387d2061f56fefb470892f140d0dc511bd4d9bb61
│   ├─5501 mysqld
│   └─6404 /bin/bash
```

Now we know a little about `control groups` mechanism, how to use it manually and what's purpose of this mechanism. Time to look inside of the Linux kernel source code and start to dive into implementation of this mechanism.

Early initialization of control groups

Now after we just saw little theory about `control groups` Linux kernel mechanism, we may start to dive into the source code of Linux kernel to acquainted with this mechanism closer. As always we will start from the initialization of `control groups`. Initialization of `cgroups` divided into two parts in the Linux kernel: early and late. In this part we will consider only `early` part and `late` part will be considered in next parts.

Early initialization of `cgroups` starts from the call of the:

```
cgroup_init_early();
```

function in the init/main.c during early initialization of the Linux kernel. This function is defined in the kernel/cgroup.c source code file and starts from the definition of two following local variables:

```
int __init cgroup_init_early(void)
{
    static struct cgroup_sb_opts __initdata opts;
    struct cgroup_subsys *ss;
    ...
    ...
    ...
}
```

The `cgroup_sb_opts` structure defined in the same source code file and looks:

```
struct cgroup_sb_opts {
    u16 subsys_mask;
    unsigned int flags;
    char *release_agent;
    bool cpuset_clone_children;
    char *name;
    bool none;
};
```

which represents mount options of `cgroupfs`. For example we may create named cgroup hierarchy (with name `my_cgrp`) with the `name=` option and without any subsystems:

```
$ mount -t cgroup -oname=my_cgrp,none /mnt/cgroups
```

The second variable - `ss` has type - `cgroup_subsys` structure which is defined in the include/linux/cgroup-defs.h header file and as you may guess from the name of the type, it represents a `cgroup` subsystem. This structure contains various fields and callback functions like:

```
struct cgroup_subsys {
    int (*css_online)(struct cgroup_subsys_state *css);
    void (*css_offline)(struct cgroup_subsys_state *css);
    ...
    ...
    ...
    bool early_init:1;
    int id;
    const char *name;
    struct cgroup_root *root;
    ...
    ...
    ...
}
```

Where for example `ccs_online` and `ccs_offline` callbacks are called after a cgroup successfully will complete all allocations and a cgroup will be before releasing respectively. The `early_init` flags marks subsystems which may/should be initialized early. The `id` and `name` fields represents unique identifier in the array of registered subsystems for a cgroup and `name` of a subsystem respectively. The last - `root` fields represents pointer to the root of of a cgroup hierarchy.

Of course the `cgroup_subsys` structure is bigger and has other fields, but it is enough for now. Now as we got to know important structures related to `cgroups` mechanism, let's return to the `cgroup_init_early` function. Main purpose of this function is to do early initialization of some subsystems. As you already may guess, these `early` subsystems should have `cgroup_subsys->early_init = 1`. Let's look what subsystems may be initialized early.

After the definition of the two local variables we may see following lines of code:

```
init_cgroup_root(&cgrp_dfl_root, &opts);
cgrp_dfl_root.cgrp.self.flags |= CSS_NO_REF;
```

Here we may see call of the `init_cgroup_root` function which will execute initialization of the default unified hierarchy and after this we set `CSS_NO_REF` flag in state of this default `cgroup` to disable reference counting for this css. The `cgrp_dfl_root` is defined in the same source code file:

```
struct cgroup_root cgrp_dfl_root;
```

Its `cgrp` field represented by the `cgroup` structure which represents a `cgroup` as you already may guess and defined in the include/linux/cgroup-defs.h header file. We already know that a process which is represented by the `task_struct` in the Linux kernel. The `task_struct` does not contain direct link to a `cgroup` where this task is attached. But it may be reached via `ccs_set` field of the `task_struct`. This `ccs_set` structure holds pointer to the array of subsystem states:

```
struct css_set {
    ...
    ...
    ....
    struct cgroup_subsys_state *subsys[CGROUP_SUBSYS_COUNT];
    ...
    ...
    ...
}
```

And via the `cgroup_subsys_state`, a process may get a `cgroup` that this process is attached to:

```
struct cgroup_subsys_state {
    ...
    ...
    ...
    struct cgroup *cgroup;
    ...
    ...
    ...
}
```

So, the overall picture of `cgroups` related data structure is following:

```
+-------------+         +---------------------+    +------------->+---------------------+          +-----------
-----+
| task_struct |         |       css_set       |    |              | cgroup_subsys_state |          |     cgroup
     |
+-------------+         |                     |    |              +---------------------+          +-----------
-----+
|             |         |                     |    |              |                     |          |     flags
     |
|             |         |                     |    |              +---------------------+          |  cgroup.pr
ocs  |
|             |         |                     |    |              |       cgroup        |--------->|       id
     |
|             |         |                     |    |              +---------------------+          |      ....
     |
|-------------+         |---------------------+----+                                               +-----------
-----+
|   cgroups   | ------> | cgroup_subsys_state | array of cgroup_subsys_state
|-------------+         +---------------------+------------------>+---------------------+          +-----------
-----+
|             |         |                     |                   | cgroup_subsys_state |          |      cgrou
p    |
+-------------+         +---------------------+                   +---------------------+          +-----------
-----+
                                                                  |                     |          |      flags

     |
                                                                  +---------------------+          |   cgroup.p
rocs |
                                                                  |       cgroup        |--------->|       id

     |
                                                                  +---------------------+          |       ....

     |
                                                                  |    cgroup_subsys    |          +-----------

-----+
                                                                  +---------------------+
                                                                             |
                                                                             |
                                                                             ↓
                                                                  +---------------------+
                                                                  |    cgroup_subsys    |
                                                                  +---------------------+
                                                                  |         id          |
                                                                  |        name         |
                                                                  |     css_online      |
                                                                  |     css_ofline      |
                                                                  |       attach        |
                                                                  |        ....         |
                                                                  +---------------------+
```

So, the `init_cgroup_root` fills the `cgrp_dfl_root` with the default values. The next thing is assigning initial `ccs_set` to the `init_task` which represents first process in the system:

```
RCU_INIT_POINTER(init_task.cgroups, &init_css_set);
```

And the last big thing in the `cgroup_init_early` function is initialization of `early cgroups`. Here we go over all registered subsystems and assign unique identity number, name of a subsystem and call the `cgroup_init_subsys` function for subsystems which are marked as early:

```
for_each_subsys(ss, i) {
        ss->id = i;
        ss->name = cgroup_subsys_name[i];

        if (ss->early_init)
            cgroup_init_subsys(ss, true);
}
```

The `for_each_subsys` here is a macro which is defined in the kernel/cgroup.c source code file and just expands to the `for` loop over `cgroup_subsys` array. Definition of this array may be found in the same source code file and it looks in a little unusual way:

```
#define SUBSYS(_x) [_x ## _cgrp_id] = &_x ## _cgrp_subsys,
    static struct cgroup_subsys *cgroup_subsys[] = {
        #include <linux/cgroup_subsys.h>
};
#undef SUBSYS
```

It is defined as `SUBSYS` macro which takes one argument (name of a subsystem) and defines `cgroup_subsys` array of cgroup subsystems. Additionally we may see that the array is initialized with content of the linux/cgroup_subsys.h header file. If we will look inside of this header file we will see again set of the `SUBSYS` macros with the given subsystems names:

```
#if IS_ENABLED(CONFIG_CPUSETS)
SUBSYS(cpuset)
#endif

#if IS_ENABLED(CONFIG_CGROUP_SCHED)
SUBSYS(cpu)
#endif
...
...
...
```

This works because of `#undef` statement after first definition of the `SUBSYS` macro. Look at the `&_x ## _cgrp_subsys` expression. The `##` operator concatenates right and left expression in a `C` macro. So as we passed `cpuset`, `cpu` and etc., to the `SUBSYS` macro, somewhere `cpuset_cgrp_subsys`, `cp_cgrp_subsys` should be defined. And that's true. If you will look in the kernel/cpuset.c source code file, you will see this definition:

```
struct cgroup_subsys cpuset_cgrp_subsys = {
    ...
    ...
    ...
    .early_init    = true,
};
```

So the last step in the `cgroup_init_early` function is initialization of early subsystems with the call of the `cgroup_init_subsys` function. Following early subsystems will be initialized:

- `cpuset`;
- `cpu`;
- `cpuacct`.

The `cgroup_init_subsys` function does initialization of the given subsystem with the default values. For example sets root of hierarchy, allocates space for the given subsystem with the call of the `css_alloc` callback function, link a subsystem with a parent if it exists, add allocated subsystem to the initial process and etc.

That's all. From this moment early subsystems are initialized.

Conclusion

It is the end of the first part which describes introduction into `Control groups` mechanism in the Linux kernel. We covered some theory and the first steps of initialization of stuffs related to `control groups` mechanism. In the next part we will continue to dive into the more practical aspects of `control groups`.

If you have any questions or suggestions write me a comment or ping me at twitter.

Please note that English is not my first language, And I am really sorry for any inconvenience. If you find any mistakes please send me a PR to linux-insides.

Links

- control groups
- PID
- cpuset
- block devices
- huge pages
- sysfs
- proc
- cgroups kernel documentation
- cgroups v2
- bash
- docker)
- perf events)
- Previous chapter

Linux kernel concepts

This chapter describes various concepts which are used in the Linux kernel.

- Per-CPU variables
- CPU masks
- The initcall mechanism
- Notification Chains

Per-CPU variables

Per-CPU variables are one of the kernel features. You can understand the meaning of this feature by reading its name. We can create a variable and each processor core will have its own copy of this variable. In this part, we take a closer look at this feature and try to understand how it is implemented and how it works.

The kernel provides an API for creating per-cpu variables - the `DEFINE_PER_CPU` macro:

```
#define DEFINE_PER_CPU(type, name) \
        DEFINE_PER_CPU_SECTION(type, name, "")
```

This macro defined in the include/linux/percpu-defs.h as many other macros for work with per-cpu variables. Now we will see how this feature is implemented.

Take a look at the `DEFINE_PER_CPU` definition. We see that it takes 2 parameters: `type` and `name` , so we can use it to create per-cpu variables, for example like this:

```
DEFINE_PER_CPU(int, per_cpu_n)
```

We pass the type and the name of our variable. `DEFINE_PER_CPU` calls the `DEFINE_PER_CPU_SECTION` macro and passes the same two parameters and empty string to it. Let's look at the definition of the `DEFINE_PER_CPU_SECTION` :

```
#define DEFINE_PER_CPU_SECTION(type, name, sec)    \
         __PCPU_ATTRS(sec) PER_CPU_DEF_ATTRIBUTES  \
         __typeof__(type) name
```

```
#define __PCPU_ATTRS(sec)                                                \
         __percpu __attribute__((section(PER_CPU_BASE_SECTION sec)))     \
         PER_CPU_ATTRIBUTES
```

where `section` is:

```
#define PER_CPU_BASE_SECTION ".data..percpu"
```

After all macros are expanded we will get a global per-cpu variable:

```
__attribute__((section(".data..percpu"))) int per_cpu_n
```

It means that we will have a `per_cpu_n` variable in the `.data..percpu` section. We can find this section in the `vmlinux` :

```
.data..percpu 00013a58  0000000000000000  000000001a5c000  00e00000  2**12
              CONTENTS, ALLOC, LOAD, DATA
```

Ok, now we know that when we use the `DEFINE_PER_CPU` macro, a per-cpu variable in the `.data..percpu` section will be created. When the kernel initializes it calls the `setup_per_cpu_areas` function which loads the `.data..percpu` section multiple times, one section per CPU.

Let's look at the per-CPU areas initialization process. It starts in the init/main.c from the call of the `setup_per_cpu_areas` function which is defined in the arch/x86/kernel/setup_percpu.c.

```
pr_info("NR_CPUS:%d nr_cpumask_bits:%d nr_cpu_ids:%d nr_node_ids:%d\n",
```

```
        NR_CPUS, nr_cpumask_bits, nr_cpu_ids, nr_node_ids);
```

The `setup_per_cpu_areas` starts from the output information about the maximum number of CPUs set during kernel configuration with the `CONFIG_NR_CPUS` configuration option, actual number of CPUs, `nr_cpumask_bits` is the same that `NR_CPUS` bit for the new `cpumask` operators and number of `NUMA` nodes.

We can see this output in the dmesg:

```
$ dmesg | grep percpu
[    0.000000] setup_percpu: NR_CPUS:8 nr_cpumask_bits:8 nr_cpu_ids:8 nr_node_ids:1
```

In the next step we check the `percpu` first chunk allocator. All percpu areas are allocated in chunks. The first chunk is used for the static percpu variables. The Linux kernel has `percpu_alloc` command line parameters which provides the type of the first chunk allocator. We can read about it in the kernel documentation:

```
percpu_alloc=    Select which percpu first chunk allocator to use.
        Currently supported values are "embed" and "page".
        Archs may support subset or none of the    selections.
        See comments in mm/percpu.c for details on each
        allocator.  This parameter is primarily    for debugging
        and performance comparison.
```

The mm/percpu.c contains the handler of this command line option:

```
early_param("percpu_alloc", percpu_alloc_setup);
```

Where the `percpu_alloc_setup` function sets the `pcpu_chosen_fc` variable depends on the `percpu_alloc` parameter value. By default the first chunk allocator is `auto` :

```
enum pcpu_fc pcpu_chosen_fc __initdata = PCPU_FC_AUTO;
```

If the `percpu_alloc` parameter is not given to the kernel command line, the `embed` allocator will be used which embeds the first percpu chunk into bootmem with the memblock. The last allocator is the first chunk `page` allocator which maps the first chunk with `PAGE_SIZE` pages.

As I wrote above, first of all we make a check of the first chunk allocator type in the `setup_per_cpu_areas` . We check that first chunk allocator is not page:

```
if (pcpu_chosen_fc != PCPU_FC_PAGE) {
    ...
    ...
    ...
}
```

If it is not `PCPU_FC_PAGE` , we will use the `embed` allocator and allocate space for the first chunk with the `pcpu_embed_first_chunk` function:

```
rc = pcpu_embed_first_chunk(PERCPU_FIRST_CHUNK_RESERVE,
                        dyn_size, atom_size,
                        pcpu_cpu_distance,
                        pcpu_fc_alloc, pcpu_fc_free);
```

As shown above, the `pcpu_embed_first_chunk` function embeds the first percpu chunk into bootmem then we pass a couple of parameters to the `pcup_embed_first_chunk` . They are as follows:

* `PERCPU_FIRST_CHUNK_RESERVE` - the size of the reserved space for the static `percpu` variables;
* `dyn_size` - minimum free size for dynamic allocation in bytes;
* `atom_size` - all allocations are whole multiples of this and aligned to this parameter;
* `pcpu_cpu_distance` - callback to determine distance between cpus;
* `pcpu_fc_alloc` - function to allocate `percpu` page;
* `pcpu_fc_free` - function to release `percpu` page.

We calculate all of these parameters before the call of the `pcpu_embed_first_chunk` :

```
const size_t dyn_size = PERCPU_MODULE_RESERVE + PERCPU_DYNAMIC_RESERVE - PERCPU_FIRST_CHUNK_RESERVE;
size_t atom_size;
#ifdef CONFIG_X86_64
        atom_size = PMD_SIZE;
#else
        atom_size = PAGE_SIZE;
#endif
```

If the first chunk allocator is `PCPU_FC_PAGE` , we will use the `pcpu_page_first_chunk` instead of the `pcpu_embed_first_chunk` . After that `percpu` areas up, we setup `percpu` offset and its segment for every CPU with the `setup_percpu_segment` function (only for `x86` systems) and move some early data from the arrays to the `percpu` variables (`x86_cpu_to_apicid` , `irq_stack_ptr` and etc...). After the kernel finishes the initialization process, we will have loaded N `.data..percpu` sections, where N is the number of CPUs, and the section used by the bootstrap processor will contain an uninitialized variable created with the `DEFINE_PER_CPU` macro.

The kernel provides an API for per-cpu variables manipulating:

* get_cpu_var(var)
* put_cpu_var(var)

Let's look at the `get_cpu_var` implementation:

```
#define get_cpu_var(var)     \
(*({                         \
         preempt_disable();  \
         this_cpu_ptr(&var); \
}))
```

The Linux kernel is preemptible and accessing a per-cpu variable requires us to know which processor the kernel is running on. So, current code must not be preempted and moved to the another CPU while accessing a per-cpu variable. That's why, first of all we can see a call of the `preempt_disable` function then a call of the `this_cpu_ptr` macro, which looks like:

```
#define this_cpu_ptr(ptr) raw_cpu_ptr(ptr)
```

and

```
#define raw_cpu_ptr(ptr)        per_cpu_ptr(ptr, 0)
```

where `per_cpu_ptr` returns a pointer to the per-cpu variable for the given cpu (second parameter). After we've created a per-cpu variable and made modifications to it, we must call the `put_cpu_var` macro which enables preemption with a call of `preempt_enable` function. So the typical usage of a per-cpu variable is as follows:

```
get_cpu_var(var);
...
//Do something with the 'var'
...
put_cpu_var(var);
```

Let's look at the `per_cpu_ptr` macro:

```
#define per_cpu_ptr(ptr, cpu)                             \
({                                                        \
        __verify_pcpu_ptr(ptr);                           \
         SHIFT_PERCPU_PTR((ptr), per_cpu_offset((cpu)));  \
})
```

As I wrote above, this macro returns a per-cpu variable for the given cpu. First of all it calls `__verify_pcpu_ptr` :

```
#define __verify_pcpu_ptr(ptr)
do {
    const void __percpu *__vpp_verify = (typeof((ptr) + 0))NULL;
    (void)__vpp_verify;
} while (0)
```

which makes the given `ptr` type of `const void __percpu *` ,

After this we can see the call of the `SHIFT_PERCPU_PTR` macro with two parameters. As first parameter we pass our ptr and for second parameter we pass the cpu number to the `per_cpu_offset` macro:

```
#define per_cpu_offset(x) (__per_cpu_offset[x])
```

which expands to getting the `x` element from the `__per_cpu_offset` array:

```
extern unsigned long __per_cpu_offset[NR_CPUS];
```

where `NR_CPUS` is the number of CPUs. The `__per_cpu_offset` array is filled with the distances between cpu-variable copies. For example all per-cpu data is `X` bytes in size, so if we access `__per_cpu_offset[Y]` , `X*Y` will be accessed. Let's look at the `SHIFT_PERCPU_PTR` implementation:

```
#define SHIFT_PERCPU_PTR(__p, __offset)                                 \
         RELOC_HIDE((typeof(*(__p)) __kernel __force *)(__p), (__offset))
```

`RELOC_HIDE` just returns offset `(typeof(ptr)) (__ptr + (off))` and it will return a pointer to the variable.

That's all! Of course it is not the full API, but a general overview. It can be hard to start with, but to understand per-cpu variables you mainly need to understand the include/linux/percpu-defs.h magic.

Let's again look at the algorithm of getting a pointer to a per-cpu variable:

- The kernel creates multiple `.data..percpu` sections (one per-cpu) during initialization process;
- All variables created with the `DEFINE_PER_CPU` macro will be relocated to the first section or for CPU0;
- `__per_cpu_offset` array filled with the distance (`BOOT_PERCPU_OFFSET`) between `.data..percpu` sections;
- When the `per_cpu_ptr` is called, for example for getting a pointer on a certain per-cpu variable for the third CPU, the `__per_cpu_offset` array will be accessed, where every index points to the required CPU.

That's all.

CPU masks

Introduction

`Cpumasks` is a special way provided by the Linux kernel to store information about CPUs in the system. The relevant source code and header files which contains API for `Cpumasks` manipulation:

* include/linux/cpumask.h
* lib/cpumask.c
* kernel/cpu.c

As comment says from the include/linux/cpumask.h: Cpumasks provide a bitmap suitable for representing the set of CPU's in a system, one bit position per CPU number. We already saw a bit about cpumask in the `boot_cpu_init` function from the Kernel entry point part. This function makes first boot cpu online, active and etc...:

```
set_cpu_online(cpu, true);
set_cpu_active(cpu, true);
set_cpu_present(cpu, true);
set_cpu_possible(cpu, true);
```

Before we will consider implementation of these functions, let's consider all of these masks.

The `cpu_possible` is a set of cpu ID's which can be plugged in anytime during the life of that system boot or in other words mask of possible CPUs contains maximum number of CPUs which are possible in the system. It will be equal to value of the `NR_CPUS` which is which is set statically via the `CONFIG_NR_CPUS` kernel configuration option.

The `cpu_present` mask represents which CPUs are currently plugged in.

The `cpu_online` represents a subset of the `cpu_present` and indicates CPUs which are available for scheduling or in other words a bit from this mask tells to kernel is a processor may be utilized by the Linux kernel.

The last mask is `cpu_active` . Bits of this mask tells to Linux kernel is a task may be moved to a certain processor.

All of these masks depend on the `CONFIG_HOTPLUG_CPU` configuration option and if this option is disabled `possible == present` and `active == online` . The implementations of all of these functions are very similar. Every function checks the second parameter. If it is `true` , it calls `cpumask_set_cpu` otherwise it calls `cpumask_clear_cpu` .

There are two ways for a `cpumask` creation. First is to use `cpumask_t` . It is defined as:

```
typedef struct cpumask { DECLARE_BITMAP(bits, NR_CPUS); } cpumask_t;
```

It wraps the `cpumask` structure which contains one bitmask `bits` field. The `DECLARE_BITMAP` macro gets two parameters:

* bitmap name;
* number of bits.

and creates an array of `unsigned long` with the given name. Its implementation is pretty easy:

```
#define DECLARE_BITMAP(name,bits) \
        unsigned long name[BITS_TO_LONGS(bits)]
```

where `BITS_TO_LONGS` :

```
#define BITS_TO_LONGS(nr)       DIV_ROUND_UP(nr, BITS_PER_BYTE * sizeof(long))
```

```
#define DIV_ROUND_UP(n,d) (((n) + (d) - 1) / (d))
```

As we are focusing on the `x86_64` architecture, `unsigned long` is 8-bytes size and our array will contain only one element:

```
(((8) + (8) - 1) / (8)) = 1
```

`NR_CPUS` macro represents the number of CPUs in the system and depends on the `CONFIG_NR_CPUS` macro which is defined in include/linux/threads.h and looks like this:

```
#ifndef CONFIG_NR_CPUS
        #define CONFIG_NR_CPUS  1
#endif

#define NR_CPUS         CONFIG_NR_CPUS
```

The second way to define cpumask is to use the `DECLARE_BITMAP` macro directly and the `to_cpumask` macro which converts the given bitmap to `struct cpumask *` :

```
#define to_cpumask(bitmap)                                              \
        ((struct cpumask *)(1 ? (bitmap)                                \
                            : (void *)sizeof(__check_is_bitmap(bitmap))))
```

We can see the ternary operator operator here which is `true` every time. `__check_is_bitmap` inline function is defined as:

```
static inline int __check_is_bitmap(const unsigned long *bitmap)
{
        return 1;
}
```

And returns `1` every time. We need it here for only one purpose: at compile time it checks that a given `bitmap` is a bitmap, or in other words it checks that a given `bitmap` has type - `unsigned long *` . So we just pass `cpu_possible_bits` to the `to_cpumask` macro for converting an array of `unsigned long` to the `struct cpumask *` .

cpumask API

As we can define cpumask with one of the method, Linux kernel provides API for manipulating a cpumask. Let's consider one of the function which presented above. For example `set_cpu_online` . This function takes two parameters:

- Number of CPU;
- CPU status;

Implementation of this function looks as:

```
void set_cpu_online(unsigned int cpu, bool online)
{
    if (online) {
        cpumask_set_cpu(cpu, to_cpumask(cpu_online_bits));
        cpumask_set_cpu(cpu, to_cpumask(cpu_active_bits));
    } else {
        cpumask_clear_cpu(cpu, to_cpumask(cpu_online_bits));
    }
}
```

First of all it checks the second `state` parameter and calls `cpumask_set_cpu` or `cpumask_clear_cpu` depends on it. Here we can see casting to the `struct cpumask *` of the second parameter in the `cpumask_set_cpu`. In our case it is `cpu_online_bits` which is a bitmap and defined as:

```
static DECLARE_BITMAP(cpu_online_bits, CONFIG_NR_CPUS) __read_mostly;
```

The `cpumask_set_cpu` function makes only one call to the `set_bit` function:

```
static inline void cpumask_set_cpu(unsigned int cpu, struct cpumask *dstp)
{
        set_bit(cpumask_check(cpu), cpumask_bits(dstp));
}
```

The `set_bit` function takes two parameters too, and sets a given bit (first parameter) in the memory (second parameter or `cpu_online_bits` bitmap). We can see here that before `set_bit` will be called, its two parameters will be passed to the

* cpumask_check;
* cpumask_bits.

Let's consider these two macros. First if `cpumask_check` does nothing in our case and just returns given parameter. The second `cpumask_bits` just returns the `bits` field from the given `struct cpumask *` structure:

```
#define cpumask_bits(maskp) ((maskp)->bits)
```

Now let's look on the `set_bit` implementation:

```
 static __always_inline void
 set_bit(long nr, volatile unsigned long *addr)
 {
         if (IS_IMMEDIATE(nr)) {
                asm volatile(LOCK_PREFIX "orb %1,%0"
                        : CONST_MASK_ADDR(nr, addr)
                        : "iq" ((u8)CONST_MASK(nr))
                        : "memory");
        } else {
                asm volatile(LOCK_PREFIX "bts %1,%0"
                        : BITOP_ADDR(addr) : "Ir" (nr) : "memory");
        }
 }
```

This function looks scary, but it is not so hard as it seems. First of all it passes `nr` or number of the bit to the `IS_IMMEDIATE` macro which just calls the GCC internal `__builtin_constant_p` function:

```
#define IS_IMMEDIATE(nr)    (__builtin_constant_p(nr))
```

`__builtin_constant_p` checks that given parameter is known constant at compile-time. As our `cpu` is not compile-time constant, the `else` clause will be executed:

```
asm volatile(LOCK_PREFIX "bts %1,%0" : BITOP_ADDR(addr) : "Ir" (nr) : "memory");
```

Let's try to understand how it works step by step:

`LOCK_PREFIX` is a x86 `lock` instruction. This instruction tells the cpu to occupy the system bus while the instruction(s) will be executed. This allows the CPU to synchronize memory access, preventing simultaneous access of multiple processors (or devices - the DMA controller for example) to one memory cell.

`BITOP_ADDR` casts the given parameter to the `(*(volatile long *)` and adds `+m` constraints. `+` means that this operand is both read and written by the instruction. `m` shows that this is a memory operand. `BITOP_ADDR` is defined as:

```
#define BITOP_ADDR(x) "+m" (*(volatile long *) (x))
```

Next is the `memory` clobber. It tells the compiler that the assembly code performs memory reads or writes to items other than those listed in the input and output operands (for example, accessing the memory pointed to by one of the input parameters).

`Ir` - immediate register operand.

The `bts` instruction sets a given bit in a bit string and stores the value of a given bit in the `CF` flag. So we passed the cpu number which is zero in our case and after `set_bit` is executed, it sets the zero bit in the `cpu_online_bits` cpumask. It means that the first cpu is online at this moment.

Besides the `set_cpu_*` API, cpumask of course provides another API for cpumasks manipulation. Let's consider it in short.

Additional cpumask API

cpumask provides a set of macros for getting the numbers of CPUs in various states. For example:

```
#define num_online_cpus()    cpumask_weight(cpu_online_mask)
```

This macro returns the amount of `online` CPUs. It calls the `cpumask_weight` function with the `cpu_online_mask` bitmap (read about it). The `cpumask_weight` function makes one call of the `bitmap_weight` function with two parameters:

- cpumask bitmap;
- `nr_cpumask_bits` - which is `NR_CPUS` in our case.

```
static inline unsigned int cpumask_weight(const struct cpumask *srcp)
{
    return bitmap_weight(cpumask_bits(srcp), nr_cpumask_bits);
}
```

and calculates the number of bits in the given bitmap. Besides the `num_online_cpus`, cpumask provides macros for the all CPU states:

- num_possible_cpus;
- num_active_cpus;
- cpu_online;
- cpu_possible.

and many more.

Besides that the Linux kernel provides the following API for the manipulation of `cpumask`:

- `for_each_cpu` - iterates over every cpu in a mask;
- `for_each_cpu_not` - iterates over every cpu in a complemented mask;
- `cpumask_clear_cpu` - clears a cpu in a cpumask;
- `cpumask_test_cpu` - tests a cpu in a mask;
- `cpumask_setall` - set all cpus in a mask;
- `cpumask_size` - returns size to allocate for a 'struct cpumask' in bytes;

and many many more...

Links

- cpumask documentation

The initcall mechanism

Introduction

As you may understand from the title, this part will cover an interesting and important concept in the Linux kernel which is called - `initcall` . We already saw definitions like these:

```
early_param("debug", debug_kernel);
```

or

```
arch_initcall(init_pit_clocksource);
```

in some parts of the Linux kernel. Before we will see how this mechanism is implemented in the Linux kernel, we must know actually what is it and how the Linux kernel uses it. Definitions like these represent a callback function which is will be called during initialization of the Linux kernel of right after. Actually the main point of the `initcall` mechanism is to determine correct order of the built-in modules and subsystems initialization. For example let's look at the following function:

```
static int __init nmi_warning_debugfs(void)
{
    debugfs_create_u64("nmi_longest_ns", 0644,
                       arch_debugfs_dir, &nmi_longest_ns);
    return 0;
}
```

from the arch/x86/kernel/nmi.c source code file. As we may see it just creates the `nmi_longest_ns` debugfs file in the `arch_debugfs_dir` directory. Actually, this `debugfs` file may be created only after the `arch_debugfs_dir` will be created. Creation of this directory occurs during the architecture-specific initialization of the Linux kernel. Actually this directory will be created in the `arch_kdebugfs_init` function from the arch/x86/kernel/kdebugfs.c source code file. Note that the `arch_kdebugfs_init` function is marked as `initcall` too:

```
arch_initcall(arch_kdebugfs_init);
```

The Linux kernel calls all architecture-specific `initcalls` before the `fs` related `initcalls` . So, our `nmi_longest_ns` file will be created only after the `arch_kdebugfs_dir` directory will be created. Actually, the Linux kernel provides eight levels of main `initcalls` :

- `early` ;
- `core` ;
- `postcore` ;
- `arch` ;
- `subsys` ;
- `fs` ;
- `device` ;
- `late` .

All of their names are represented by the `initcall_level_names` array which is defined in the init/main.c source code file:

```
static char *initcall_level_names[] __initdata = {
    "early",
    "core",
```

```
    "postcore",
    "arch",
    "subsys",
    "fs",
    "device",
    "late",
};
```

All functions which are marked as `initcall` by these identifiers, will be called in the same order or at first `early initcalls` will be called, at second `core initcalls` and etc. From this moment we know a little about `initcall` mechanism, so we can start to dive into the source code of the Linux kernel to see how this mechanism is implemented.

Implementation initcall mechanism in the Linux kernel

The Linux kernel provides a set of macros from the include/linux/init.h header file to mark a given function as `initcall`. All of these macros are pretty simple:

```
#define early_initcall(fn)        __define_initcall(fn, early)
#define core_initcall(fn)        __define_initcall(fn, 1)
#define postcore_initcall(fn)        __define_initcall(fn, 2)
#define arch_initcall(fn)        __define_initcall(fn, 3)
#define subsys_initcall(fn)        __define_initcall(fn, 4)
#define fs_initcall(fn)        __define_initcall(fn, 5)
#define device_initcall(fn)        __define_initcall(fn, 6)
#define late_initcall(fn)        __define_initcall(fn, 7)
```

and as we may see these macros just expand to the call of the `__define_initcall` macro from the same header file. Moreover, the `__define_initcall` macro takes two arguments:

- `fn` - callback function which will be called during call of `initcalls` of the certain level;
- `id` - identifier to identify `initcall` to prevent error when two the same `initcalls` point to the same handler.

The implementation of the `__define_initcall` macro looks like:

```
#define __define_initcall(fn, id) \
    static initcall_t __initcall_##fn##id __used \
    __attribute__((__section__(".initcall" #id ".init"))) = fn; \
    LTO_REFERENCE_INITCALL(__initcall_##fn##id)
```

To understand the `__define_initcall` macro, first of all let's look at the `initcall_t` type. This type is defined in the same header file and it represents pointer to a function which returns pointer to integer which will be result of the `initcall`:

```
typedef int (*initcall_t)(void);
```

Now let's return to the `_-define_initcall` macro. The ## provides ability to concatenate two symbols. In our case, the first line of the `__define_initcall` macro produces definition of the given function which is located in the `.initcall id .init` ELF section and marked with the following gcc attributes: `__initcall_function_name_id` and `__used`. If we will look in the include/asm-generic/vmlinux.lds.h header file which represents data for the kernel linker script, we will see that all of `initcalls` sections will be placed in the `.data` section:

```
#define INIT_CALLS                    \
        VMLINUX_SYMBOL(__initcall_start) = .;    \
        *(.initcallearly.init)                    \
        INIT_CALLS_LEVEL(0)                        \
        INIT_CALLS_LEVEL(1)                        \
        INIT_CALLS_LEVEL(2)                        \
        INIT_CALLS_LEVEL(3)                        \
```

```
        INIT_CALLS_LEVEL(4)                        \
        INIT_CALLS_LEVEL(5)                        \
        INIT_CALLS_LEVEL(rootfs)                \
        INIT_CALLS_LEVEL(6)                        \
        INIT_CALLS_LEVEL(7)                        \
        VMLINUX_SYMBOL(__initcall_end) = .;

#define INIT_DATA_SECTION(initsetup_align)    \
    .init.data : AT(ADDR(.init.data) - LOAD_OFFSET) {       \
        ...                                               \
        INIT_CALLS                                          \
        ...                                               \
    }
```

The second attribute - `__used` is defined in the include/linux/compiler-gcc.h header file and it expands to the definition of the following `gcc` attribute:

```
#define __used   __attribute__((__used__))
```

which prevents `variable defined but not used` warning. The last line of the `__define_initcall` macro is:

```
LTO_REFERENCE_INITCALL(__initcall_##fn##id)
```

depends on the `CONFIG_LTO` kernel configuration option and just provides stub for the compiler Link time optimization:

```
#ifdef CONFIG_LTO
#define LTO_REFERENCE_INITCALL(x) \
        static __used __exit void *reference_##x(void)  \
        {                                               \
                return &x;                              \
        }
#else
#define LTO_REFERENCE_INITCALL(x)
#endif
```

In order to prevent any problem when there is no reference to a variable in a module, it will be moved to the end of the program. That's all about the `__define_initcall` macro. So, all of the `*_initcall` macros will be expanded during compilation of the Linux kernel, and all `initcalls` will be placed in their sections and all of them will be available from the `.data` section and the Linux kernel will know where to find a certain `initcall` to call it during initialization process.

As `initcalls` can be called by the Linux kernel, let's look how the Linux kernel does this. This process starts in the `do_basic_setup` function from the init/main.c source code file:

```
static void __init do_basic_setup(void)
{
    ...
    ...
    ...
       do_initcalls();
    ...
    ...
    ...
}
```

which is called during the initialization of the Linux kernel, right after main steps of initialization like memory manager related initialization, `CPU` subsystem and other already finished. The `do_initcalls` function just goes through the array of `initcall` levels and call the `do_initcall_level` function for each level:

```
static void __init do_initcalls(void)
```

```
{
    int level;

    for (level = 0; level < ARRAY_SIZE(initcall_levels) - 1; level++)
        do_initcall_level(level);
}
```

The `initcall_levels` array is defined in the same source code file and contains pointers to the sections which were defined in the `__define_initcall` macro:

```
static initcall_t *initcall_levels[] __initdata = {
    __initcall0_start,
    __initcall1_start,
    __initcall2_start,
    __initcall3_start,
    __initcall4_start,
    __initcall5_start,
    __initcall6_start,
    __initcall7_start,
    __initcall_end,
};
```

If you are interested, you can find these sections in the `arch/x86/kernel/vmlinux.lds` linker script which is generated after the Linux kernel compilation:

```
.init.data : AT(ADDR(.init.data) - 0xffffffff80000000) {
    ...
    ...
    ...
    ...
    __initcall_start = .;
    *(.initcallearly.init)
    __initcall0_start = .;
    *(.initcall0.init)
    *(.initcall0s.init)
    __initcall1_start = .;
    ...
    ...
}
```

If you are not familiar with this then you can know more about linkers in the special part of this book.

As we just saw, the `do_initcall_level` function takes one parameter - level of `initcall` and does following two things: First of all this function parses the `initcall_command_line` which is copy of usual kernel command line which may contain parameters for modules with the `parse_args` function from the kernel/params.c source code file and call the `do_on_initcall` function for each level:

```
for (fn = initcall_levels[level]; fn < initcall_levels[level+1]; fn++)
        do_one_initcall(*fn);
```

The `do_on_initcall` does main job for us. As we may see, this function takes one parameter which represent `initcall` callback function and does the call of the given callback:

```
int __init_or_module do_one_initcall(initcall_t fn)
{
    int count = preempt_count();
    int ret;
    char msgbuf[64];

    if (initcall_blacklisted(fn))
        return -EPERM;
```

```
    if (initcall_debug)
        ret = do_one_initcall_debug(fn);
    else
        ret = fn();

    msgbuf[0] = 0;

    if (preempt_count() != count) {
        sprintf(msgbuf, "preemption imbalance ");
        preempt_count_set(count);
    }
    if (irqs_disabled()) {
        strlcat(msgbuf, "disabled interrupts ", sizeof(msgbuf));
        local_irq_enable();
    }
    WARN(msgbuf[0], "initcall %pF returned with %s\n", fn, msgbuf);

    return ret;
}
```

Let's try to understand what does the `do_on_initcall` function does. First of all we increase preemption counter so that we can check it later to be sure that it is not imbalanced. After this step we can see the call of the `initcall_backlist` function which goes over the `blacklisted_initcalls` list which stores blacklisted `initcalls` and releases the given `initcall` if it is located in this list:

```
list_for_each_entry(entry, &blacklisted_initcalls, next) {
    if (!strcmp(fn_name, entry->buf)) {
        pr_debug("initcall %s blacklisted\n", fn_name);
        kfree(fn_name);
        return true;
    }
}
```

The blacklisted `initcalls` stored in the `blacklisted_initcalls` list and this list is filled during early Linux kernel initialization from the Linux kernel command line.

After the blacklisted `initcalls` will be handled, the next part of code does directly the call of the `initcall` :

```
if (initcall_debug)
    ret = do_one_initcall_debug(fn);
else
    ret = fn();
```

Depends on the value of the `initcall_debug` variable, the `do_one_initcall_debug` function will call `initcall` or this function will do it directly via `fn()` . The `initcall_debug` variable is defined in the same source code file:

```
bool initcall_debug;
```

and provides ability to print some information to the kernel log buffer. The value of the variable can be set from the kernel commands via the `initcall_debug` parameter. As we can read from the documentation of the Linux kernel command line:

```
initcall_debug    [KNL] Trace initcalls as they are executed.  Useful
                      for working out where the kernel is dying during
                      startup.
```

And that's true. If we will look at the implementation of the `do_one_initcall_debug` function, we will see that it does the same as the `do_one_initcall` function or i.e. the `do_one_initcall_debug` function calls the given `initcall` and prints some information (like the pid of the currently running task, duration of execution of the `initcall` and etc.) related to the execution of the given `initcall` :

```
static int __init_or_module do_one_initcall_debug(initcall_t fn)
{
    ktime_t calltime, delta, rettime;
    unsigned long long duration;
    int ret;

    printk(KERN_DEBUG "calling  %pF @ %i\n", fn, task_pid_nr(current));
    calltime = ktime_get();
    ret = fn();
    rettime = ktime_get();
    delta = ktime_sub(rettime, calltime);
    duration = (unsigned long long) ktime_to_ns(delta) >> 10;
    printk(KERN_DEBUG "initcall %pF returned %d after %lld usecs\n",
         fn, ret, duration);

    return ret;
}
```

As an `initcall` was called by the one of the `do_one_initcall` or `do_one_initcall_debug` functions, we may see two checks in the end of the `do_one_initcall` function. The first one checks the amount of possible `__preempt_count_add` and `__preempt_count_sub` calls inside of the executed initcall, and if this value is not equal to the previous value of the preemptible counter, we add the `preemption imbalance` string to the message buffer and set correct value of the preemptible counter:

```
if (preempt_count() != count) {
    sprintf(msgbuf, "preemption imbalance ");
    preempt_count_set(count);
}
```

Later this error string will be printed. The last check the state of local IRQs and if they are disabled, we add the `disabled interrupts` strings to the our message buffer and enable `IRQs` for the current processor to prevent the state when `IRQs` were disabled by an `initcall` and didn't enable again:

```
if (irqs_disabled()) {
    strlcat(msgbuf, "disabled interrupts ", sizeof(msgbuf));
    local_irq_enable();
}
```

That's all. In this way the Linux kernel does initialization of many subsystems in a correct order. From now on, we know what is the `initcall` mechanism in the Linux kernel. In this part, we covered main general portion of the `initcall` mechanism but we left some important concepts. Let's make a short look at these concepts.

First of all, we have missed one level of `initcalls` , this is `rootfs initcalls` . You can find definition of the `rootfs_initcall` in the include/linux/init.h header file along with all similar macros which we saw in this part:

```
#define rootfs_initcall(fn)        __define_initcall(fn, rootfs)
```

As we may understand from the macro's name, its main purpose is to store callbacks which are related to the rootfs. Besides this goal, it may be useful to initialize other stuffs after initialization related to filesystems level only if devices related stuff are not initialized. For example, the decompression of the initramfs which occurred in the `populate_rootfs` function from the init/initramfs.c source code file:

```
rootfs_initcall(populate_rootfs);
```

From this place, we may see familiar output:

```
[    0.199960] Unpacking initramfs...
```

Besides the `rootfs_initcall` level, there are additional `console_initcall` , `security_initcall` and other secondary `initcall` levels. The last thing that we have missed is the set of the `*_initcall_sync` levels. Almost each `*_initcall` macro that we have seen in this part, has macro companion with the `_sync` prefix:

```
#define core_initcall_sync(fn)          __define_initcall(fn, 1s)
#define postcore_initcall_sync(fn)      __define_initcall(fn, 2s)
#define arch_initcall_sync(fn)          __define_initcall(fn, 3s)
#define subsys_initcall_sync(fn)      __define_initcall(fn, 4s)
#define fs_initcall_sync(fn)          __define_initcall(fn, 5s)
#define device_initcall_sync(fn)      __define_initcall(fn, 6s)
#define late_initcall_sync(fn)          __define_initcall(fn, 7s)
```

The main goal of these additional levels is to wait for completion of all a module related initialization routines for a certain level.

That's all.

Conclusion

In this part we saw the important mechanism of the Linux kernel which allows to call a function which depends on the current state of the Linux kernel during its initialization.

If you have questions or suggestions, feel free to ping me in twitter 0xAX, drop me email or just create issue.

Please note that English is not my first language and I am really sorry for any inconvenience. If you found any mistakes please send me PR to linux-insides..

Links

- callback
- debugfs
- integer type
- symbols concatenation
- GCC
- Link time optimization
- Introduction to linkers
- Linux kernel command line
- Process identifier
- IRQs
- rootfs
- previous part

Notification Chains in Linux Kernel

Introduction

The Linux kernel is huge piece of C) code which consists from many different subsystems. Each subsystem has its own purpose which is independent of other subsystems. But often one subsystem wants to know something from other subsystem(s). There is special mechanism in the Linux kernel which allows to solve this problem partly. The name of this mechanism is - `notification chains` and its main purpose to provide a way for different subsystems to subscribe on asynchronous events from other subsystems. Note that this mechanism is only for communication inside kernel, but there are other mechanisms for communication between kernel and userspace.

Before we will consider `notification chains` API and implementation of this API, let's look at `Notification chains` mechanism from theoretical side as we did it in other parts of this book. Everything which is related to `notification chains` mechanism is located in the include/linux/notifier.h header file and kernel/notifier.c source code file. So let's open them and start to dive.

Notification Chains related data structures

Let's start to consider `notification chains` mechanism from related data structures. As I wrote above, main data structures should be located in the include/linux/notifier.h header file, so the Linux kernel provides generic API which does not depend on certain architecture. In general, the `notification chains` mechanism represents a list (that's why it named `chains`) of callback) functions which are will be executed when an event will be occurred.

All of these callback functions are represented as `notifier_fn_t` type in the Linux kernel:

```
typedef    int (*notifier_fn_t)(struct notifier_block *nb, unsigned long action, void *data);
```

So we may see that it takes three following arguments:

- `nb` - is linked list of function pointers (will see it now);
- `action` - is type of an event. A notification chain may support multiple events, so we need this parameter to distinguish an event from other events;
- `data` - is storage for private information. Actually it allows to provide additional data information about an event.

Additionally we may see that `notifier_fn_t` returns an integer value. This integer value maybe one of:

- `NOTIFY_DONE` - subscriber does not interested in notification;
- `NOTIFY_OK` - notification was processed correctly;
- `NOTIFY_BAD` - something went wrong;
- `NOTIFY_STOP` - notification is done, but no further callbacks should be called for this event.

All of these results defined as macros in the include/linux/notifier.h header file:

```
#define NOTIFY_DONE        0x0000
#define NOTIFY_OK        0x0001
#define NOTIFY_BAD        (NOTIFY_STOP_MASK|0x0002)
#define NOTIFY_STOP        (NOTIFY_OK|NOTIFY_STOP_MASK)
```

Where `NOTIFY_STOP_MASK` represented by the:

```
#define NOTIFY_STOP_MASK    0x8000
```

macro and means that callbacks will not be called during next notifications.

Each part of the Linux kernel which wants to be notified on a certain event will should provide own `notifier_fn_t` callback function. Main role of the `notification chains` mechanism is to call certain callbacks when an asynchronous event occurred.

The main building block of the `notification chains` mechanism is the `notifier_block` structure:

```
struct notifier_block {
    notifier_fn_t notifier_call;
    struct notifier_block __rcu *next;
    int priority;
};
```

which is defined in the include/linux/notifier.h file. This struct contains pointer to callback function - `notifier_call`, link to the next notification callback and `priority` of a callback function as functions with higher priority are executed first.

The Linux kernel provides notification chains of four following types:

- Blocking notifier chains;
- SRCU notifier chains;
- Atomic notifier chains;
- Raw notifier chains.

Let's consider all of these types of notification chains by order:

In the first case for the `blocking notifier chains`, callbacks will be called/executed in process context. This means that the calls in a notification chain may be blocked.

The second `SRCU notifier chains` represent alternative form of `blocking notifier chains`. In the first case, blocking notifier chains uses `rw_semaphore` synchronization primitive to protect chain links. `SRCU` notifier chains run in process context too, but uses special form of RCU mechanism which is permissible to block in an read-side critical section.

In the third case for the `atomic notifier chains` runs in interrupt or atomic context and protected by spinlock synchronization primitive. The last `raw notifier chains` provides special type of notifier chains without any locking restrictions on callbacks. This means that protection rests on the shoulders of caller side. It is very useful when we want to protect our chain with very specific locking mechanism.

If we will look at the implementation of the `notifier_block` structure, we will see that it contains pointer to the `next` element from a notification chain list, but we have no head. Actually a head of such list is in separate structure depends on type of a notification chain. For example for the `blocking notifier chains`:

```
struct blocking_notifier_head {
    struct rw_semaphore rwsem;
    struct notifier_block __rcu *head;
};
```

or for `atomic notification chains`:

```
struct atomic_notifier_head {
    spinlock_t lock;
    struct notifier_block __rcu *head;
};
```

Now as we know a little about `notification chains` mechanism let's consider implementation of its API.

Notification Chains

Usually there are two sides in a publish/subscriber mechanisms. One side who wants to get notifications and other side(s) who generates these notifications. We will consider notification chains mechanism from both sides. We will consider `blocking notification chains` in this part, because of other types of notification chains are similar to it and differs mostly in protection mechanisms.

Before a notification producer is able to produce notification, first of all it should initialize head of a notification chain. For example let's consider notification chains related to kernel loadable modules. If we will look in the kernel/module.c source code file, we will see following definition:

```
static BLOCKING_NOTIFIER_HEAD(module_notify_list);
```

which defines head for loadable modules blocking notifier chain. The `BLOCKING_NOTIFIER_HEAD` macro is defined in the include/linux/notifier.h header file and expands to the following code:

```
#define BLOCKING_INIT_NOTIFIER_HEAD(name) do {    \
        init_rwsem(&(name)->rwsem);                          \
        (name)->head = NULL;                              \
    } while (0)
```

So we may see that it takes name of a name of a head of a blocking notifier chain and initializes read/write semaphore and set head to `NULL`. Besides the `BLOCKING_INIT_NOTIFIER_HEAD` macro, the Linux kernel additionally provides `ATOMIC_INIT_NOTIFIER_HEAD`, `RAW_INIT_NOTIFIER_HEAD` macros and `srcu_init_notifier` function for initialization atomic and other types of notification chains.

After initialization of a head of a notification chain, a subsystem which wants to receive notification from the given notification chain it should register with certain function which is depends on type of notification. If you will look in the include/linux/notifier.h header file, you will see following four function for this:

```
extern int atomic_notifier_chain_register(struct atomic_notifier_head *nh,
        struct notifier_block *nb);

extern int blocking_notifier_chain_register(struct blocking_notifier_head *nh,
        struct notifier_block *nb);

extern int raw_notifier_chain_register(struct raw_notifier_head *nh,
        struct notifier_block *nb);

extern int srcu_notifier_chain_register(struct srcu_notifier_head *nh,
        struct notifier_block *nb);
```

As I already wrote above, we will cover only blocking notification chains in the part, so let's consider implementation of the `blocking_notifier_chain_register` function. Implementation of this function is located in the kernel/notifier.c source code file and as we may see the `blocking_notifier_chain_register` takes two parameters:

- `nh` - head of a notification chain;
- `nb` - notification descriptor.

Now let's look at the implementation of the `blocking_notifier_chain_register` function:

```
int raw_notifier_chain_register(struct raw_notifier_head *nh,
        struct notifier_block *n)
{
    return notifier_chain_register(&nh->head, n);
}
```

As we may see it just returns result of the `notifier_chain_register` function from the same source code file and as we may understand this function does all job for us. Definition of the `notifier_chain_register` function looks:

```
int blocking_notifier_chain_register(struct blocking_notifier_head *nh,
        struct notifier_block *n)
{
    int ret;

    if (unlikely(system_state == SYSTEM_BOOTING))
        return notifier_chain_register(&nh->head, n);

    down_write(&nh->rwsem);
    ret = notifier_chain_register(&nh->head, n);
    up_write(&nh->rwsem);
    return ret;
}
```

As we may see implementation of the `blocking_notifier_chain_register` is pretty simple. First of all there is check which check current system state and if a system in rebooting state we just call the `notifier_chain_register`. In other way we do the same call of the `notifier_chain_register` but as you may see this call is protected with read/write semaphores. Now let's look at the implementation of the `notifier_chain_register` function:

```
static int notifier_chain_register(struct notifier_block **nl,
        struct notifier_block *n)
{
    while ((*nl) != NULL) {
        if (n->priority > (*nl)->priority)
            break;
        nl = &((*nl)->next);
    }
    n->next = *nl;
    rcu_assign_pointer(*nl, n);
    return 0;
}
```

This function just inserts new `notifier_block` (given by a subsystem which wants to get notifications) to the notification chain list. Besides subscribing on an event, subscriber may unsubscribe from a certain events with the set of `unsubscribe` functions:

```
extern int atomic_notifier_chain_unregister(struct atomic_notifier_head *nh,
        struct notifier_block *nb);

extern int blocking_notifier_chain_unregister(struct blocking_notifier_head *nh,
        struct notifier_block *nb);

extern int raw_notifier_chain_unregister(struct raw_notifier_head *nh,
        struct notifier_block *nb);

extern int srcu_notifier_chain_unregister(struct srcu_notifier_head *nh,
        struct notifier_block *nb);
```

When a producer of notifications wants to notify subscribers about an event, the `*.notifier_call_chain` function will be called. As you already may guess each type of notification chains provides own function to produce notification:

```
extern int atomic_notifier_call_chain(struct atomic_notifier_head *nh,
        unsigned long val, void *v);

extern int blocking_notifier_call_chain(struct blocking_notifier_head *nh,
        unsigned long val, void *v);

extern int raw_notifier_call_chain(struct raw_notifier_head *nh,
        unsigned long val, void *v);

extern int srcu_notifier_call_chain(struct srcu_notifier_head *nh,
        unsigned long val, void *v);
```

Let's consider implementation of the `blocking_notifier_call_chain` function. This function is defined in the kernel/notifier.c source code file:

```
int blocking_notifier_call_chain(struct blocking_notifier_head *nh,
        unsigned long val, void *v)
{
    return __blocking_notifier_call_chain(nh, val, v, -1, NULL);
}
```

and as we may see it just returns result of the `__blocking_notifier_call_chain` function. As we may see, the `blocking_notifer_call_chain` takes three parameters:

- `nh` - head of notification chain list;
- `val` - type of a notification;
- `v` - input parameter which may be used by handlers.

But the `__blocking_notifier_call_chain` function takes five parameters:

```
int __blocking_notifier_call_chain(struct blocking_notifier_head *nh,
                   unsigned long val, void *v,
                   int nr_to_call, int *nr_calls)
{
    ...
    ...
    ...
}
```

Where `nr_to_call` and `nr_calls` are number of notifier functions to be called and number of sent notifications. As you may guess the main goal of the `__blocking_notifer_call_chain` function and other functions for other notification types is to call callback function when an event occurred. Implementation of the `__blocking_notifier_call_chain` is pretty simple, it just calls the `notifier_call_chain` function from the same source code file protected with read/write semaphore:

```
int __blocking_notifier_call_chain(struct blocking_notifier_head *nh,
                   unsigned long val, void *v,
                   int nr_to_call, int *nr_calls)
{
    int ret = NOTIFY_DONE;

    if (rcu_access_pointer(nh->head)) {
        down_read(&nh->rwsem);
        ret = notifier_call_chain(&nh->head, val, v, nr_to_call,
                    nr_calls);
        up_read(&nh->rwsem);
    }
    return ret;
}
```

and returns its result. In this case all job is done by the `notifier_call_chain` function. Main purpose of this function informs registered notifiers about an asynchronous event:

```
static int notifier_call_chain(struct notifier_block **nl,
                   unsigned long val, void *v,
                   int nr_to_call, int *nr_calls)
{
    ...
    ...
    ...
    ret = nb->notifier_call(nb, val, v);
    ...
    ...
    ...
```

```
    return ret;
}
```

That's all. In generall all looks pretty simple.

Now let's consider on a simple example related to loadable modules. If we will look in the kernel/module.c. As we already saw in this part, there is:

```
static BLOCKING_NOTIFIER_HEAD(module_notify_list);
```

definition of the `module_notify_list` in the kernel/module.c source code file. This definition determines head of list of blocking notifier chains related to kernel modules. There are at least three following events:

- MODULE_STATE_LIVE
- MODULE_STATE_COMING
- MODULE_STATE_GOING

in which maybe interested some subsystems of the Linux kernel. For example tracing of kernel modules states. Instead of direct call of the `atomic_notifier_chain_register` , `blocking_notifier_chain_register` and etc., most notification chains come with a set of wrappers used to register to them. Registatrion on these modules events is going with the help of such wrapper:

```
int register_module_notifier(struct notifier_block *nb)
{
    return blocking_notifier_chain_register(&module_notify_list, nb);
}
```

If we will look in the kernel/tracepoint.c source code file, we will see such registration during initialization of tracepoints:

```
static __init int init_tracepoints(void)
{
    int ret;

    ret = register_module_notifier(&tracepoint_module_nb);
    if (ret)
        pr_warn("Failed to register tracepoint module enter notifier\n");

    return ret;
}
```

Where `tracepoint_module_nb` provides callback function:

```
static struct notifier_block tracepoint_module_nb = {
    .notifier_call = tracepoint_module_notify,
    .priority = 0,
};
```

When one of the `MODULE_STATE_LIVE` , `MODULE_STATE_COMING` or `MODULE_STATE_GOING` events occurred. For example the `MODULE_STATE_LIVE` the `MODULE_STATE_COMING` notifications will be sent during execution of the init_module system call. Or for example `MODULE_STATE_GOING` will be sent during execution of the delete_module `system call` :

```
SYSCALL_DEFINE2(delete_module, const char __user *, name_user,
        unsigned int, flags)
{
    ...
    ...
    ...
    blocking_notifier_call_chain(&module_notify_list,
                 MODULE_STATE_GOING, mod);
```

```
    ...
    ...
    ...
}
```

Thus when one of these system call will be called from userspace, the Linux kernel will send certain notification depends on a system call and the `tracepoint_module_notify` callback function will be called.

That's all.

Links

* C programming langauge)
* API
* callback)
* RCU
* spinlock
* loadable modules
* semaphore
* tracepoints
* system call
* init_module system call
* delete_module
* previous part

Data Structures in the Linux Kernel

Linux kernel provides different implementations of data structures like doubly linked list, B+ tree, priority heap and many many more.

This part considers the following data structures and algorithms:

- Doubly linked list
- Radix tree
- Bit arrays

Data Structures in the Linux Kernel

Doubly linked list

Linux kernel provides its own implementation of doubly linked list, which you can find in the include/linux/list.h. We will start `Data Structures in the Linux kernel` from the doubly linked list data structure. Why? Because it is very popular in the kernel, just try to search

First of all, let's look on the main structure in the include/linux/types.h:

```
struct list_head {
    struct list_head *next, *prev;
};
```

You can note that it is different from many implementations of doubly linked list which you have seen. For example, this doubly linked list structure from the glib library looks like :

```
struct GList {
  gpointer data;
  GList *next;
  GList *prev;
};
```

Usually a linked list structure contains a pointer to the item. The implementation of linked list in Linux kernel does not. So the main question is - `where does the list store the data?`. The actual implementation of linked list in the kernel is - `Intrusive list`. An intrusive linked list does not contain data in its nodes - A node just contains pointers to the next and previous node and list nodes part of the data that are added to the list. This makes the data structure generic, so it does not care about entry data type anymore.

For example:

```
struct nmi_desc {
    spinlock_t lock;
    struct list_head head;
};
```

Let's look at some examples to understand how `list_head` is used in the kernel. As I already wrote about, there are many, really many different places where lists are used in the kernel. Let's look for an example in miscellaneous character drivers. Misc character drivers API from the drivers/char/misc.c is used for writing small drivers for handling simple hardware or virtual devices. Those drivers share same major number:

```
#define MISC_MAJOR              10
```

but have their own minor number. For example you can see it with:

```
ls -l /dev |  grep 10
crw-------   1 root root     10, 235 Mar 21 12:01 autofs
drwxr-xr-x  10 root root         200 Mar 21 12:01 cpu
crw-------   1 root root     10,  62 Mar 21 12:01 cpu_dma_latency
crw-------   1 root root     10, 203 Mar 21 12:01 cuse
drwxr-xr-x   2 root root         100 Mar 21 12:01 dri
crw-rw-rw-   1 root root     10, 229 Mar 21 12:01 fuse
crw-------   1 root root     10, 228 Mar 21 12:01 hpet
```

```
crw-------   1 root root     10, 183 Mar 21 12:01 hwrng
crw-rw----+  1 root kvm      10, 232 Mar 21 12:01 kvm
crw-rw----   1 root disk     10, 237 Mar 21 12:01 loop-control
crw-------   1 root root     10, 227 Mar 21 12:01 mcelog
crw-------   1 root root     10,  59 Mar 21 12:01 memory_bandwidth
crw-------   1 root root     10,  61 Mar 21 12:01 network_latency
crw-------   1 root root     10,  60 Mar 21 12:01 network_throughput
crw-r-----   1 root kmem     10, 144 Mar 21 12:01 nvram
brw-rw----   1 root disk      1,  10 Mar 21 12:01 ram10
crw--w----   1 root tty       4,  10 Mar 21 12:01 tty10
crw-rw----   1 root dialout   4,  74 Mar 21 12:01 ttyS10
crw-------   1 root root     10,  63 Mar 21 12:01 vga_arbiter
crw-------   1 root root     10, 137 Mar 21 12:01 vhci
```

Now let's have a close look at how lists are used in the misc device drivers. First of all, let's look on `miscdevice` structure:

```
struct miscdevice
{
      int minor;
      const char *name;
      const struct file_operations *fops;
      struct list_head list;
      struct device *parent;
      struct device *this_device;
      const char *nodename;
      mode_t mode;
};
```

We can see the fourth field in the `miscdevice` structure - `list` which is a list of registered devices. In the beginning of the source code file we can see the definition of misc_list:

```
static LIST_HEAD(misc_list);
```

which expands to the definition of variables with `list_head` type:

```
#define LIST_HEAD(name) \
    struct list_head name = LIST_HEAD_INIT(name)
```

and initializes it with the `LIST_HEAD_INIT` macro, which sets previous and next entries with the address of variable - name:

```
#define LIST_HEAD_INIT(name) { &(name), &(name) }
```

Now let's look on the `misc_register` function which registers a miscellaneous device. At the start it initializes `miscdevice->list` with the `INIT_LIST_HEAD` function:

```
INIT_LIST_HEAD(&misc->list);
```

which does the same as the `LIST_HEAD_INIT` macro:

```
static inline void INIT_LIST_HEAD(struct list_head *list)
{
    list->next = list;
    list->prev = list;
}
```

In the next step after a device is created by the `device_create` function, we add it to the miscellaneous devices list with:

```
list_add(&misc->list, &misc_list);
```

Kernel `list.h` provides this API for the addition of a new entry to the list. Let's look at its implementation:

```
static inline void list_add(struct list_head *new, struct list_head *head)
{
    __list_add(new, head, head->next);
}
```

It just calls internal function `__list_add` with the 3 given parameters:

- new - new entry.
- head - list head after which the new item will be inserted.
- head->next - next item after list head.

Implementation of the `__list_add` is pretty simple:

```
static inline void __list_add(struct list_head *new,
                  struct list_head *prev,
                  struct list_head *next)
{
    next->prev = new;
    new->next = next;
    new->prev = prev;
    prev->next = new;
}
```

Here we add a new item between `prev` and `next`. So `misc` list which we defined at the start with the `LIST_HEAD_INIT` macro will contain previous and next pointers to the `miscdevice->list`.

There is still one question: how to get list's entry. There is a special macro:

```
#define list_entry(ptr, type, member) \
    container_of(ptr, type, member)
```

which gets three parameters:

- ptr - the structure list_head pointer;
- type - structure type;
- member - the name of the list_head within the structure;

For example:

```
const struct miscdevice *p = list_entry(v, struct miscdevice, list)
```

After this we can access to any `miscdevice` field with `p->minor` or `p->name` and etc... Let's look on the `list_entry` implementation:

```
#define list_entry(ptr, type, member) \
    container_of(ptr, type, member)
```

As we can see it just calls `container_of` macro with the same arguments. At first sight, the `container_of` looks strange:

```
#define container_of(ptr, type, member) ({                      \
    const typeof( ((type *)0)->member ) *__mptr = (ptr);    \
    (type *)( (char *)__mptr - offsetof(type,member) );})
```

First of all you can note that it consists of two expressions in curly brackets. The compiler will evaluate the whole block in the curly braces and use the value of the last expression.

For example:

```
#include <stdio.h>

int main() {
    int i = 0;
    printf("i = %d\n", ({++i; ++i;}));
    return 0;
}
```

will print `2`.

The next point is `typeof`, it's simple. As you can understand from its name, it just returns the type of the given variable. When I first saw the implementation of the `container_of` macro, the strangest thing I found was the zero in the `((type *)0)` expression. Actually this pointer magic calculates the offset of the given field from the address of the structure, but as we have `0` here, it will be just a zero offset along with the field width. Let's look at a simple example:

```
#include <stdio.h>

struct s {
        int field1;
        char field2;
        char field3;
};

int main() {
    printf("%p\n", &((struct s*)0)->field3);
    return 0;
}
```

will print `0x5`.

The next `offsetof` macro calculates offset from the beginning of the structure to the given structure's field. Its implementation is very similar to the previous code:

```
#define offsetof(TYPE, MEMBER) ((size_t) &((TYPE *)0)->MEMBER)
```

Let's summarize all about `container_of` macro. The `container_of` macro returns the address of the structure by the given address of the structure's field with `list_head` type, the name of the structure field with `list_head` type and type of the container structure. At the first line this macro declares the `__mptr` pointer which points to the field of the structure that `ptr` points to and assigns `ptr` to it. Now `ptr` and `__mptr` point to the same address. Technically we don't need this line but it's useful for type checking. The first line ensures that the given structure (`type` parameter) has a member called `member`. In the second line it calculates offset of the field from the structure with the `offsetof` macro and subtracts it from the structure address. That's all.

Of course `list_add` and `list_entry` is not the only functions which `<linux/list.h>` provides. Implementation of the doubly linked list provides the following API:

- list_add
- list_add_tail
- list_del
- list_replace
- list_move
- list_is_last

* list_empty
* list_cut_position
* list_splice
* list_for_each
* list_for_each_entry

and many more.

Data Structures in the Linux Kernel

Radix tree

As you already know linux kernel provides many different libraries and functions which implement different data structures and algorithms. In this part we will consider one of these data structures - Radix tree. There are two files which are related to `radix tree` implementation and API in the linux kernel:

- include/linux/radix-tree.h
- lib/radix-tree.c

Lets talk about what a `radix tree` is. Radix tree is a `compressed trie` where a trie is a data structure which implements an interface of an associative array and allows to store values as `key-value`. The keys are usually strings, but any data type can be used. A trie is different from an `n-tree` because of its nodes. Nodes of a trie do not store keys; instead, a node of a trie stores single character labels. The key which is related to a given node is derived by traversing from the root of the tree to this node. For example:

```
               +-----------+
               |           |
               |    " "    |
               |           |
        +------+-----------+------+
        |                         |
        |                         |
   +----v------+            +-----v-----+
   |           |            |           |
   |    g      |            |     c     |
   |           |            |           |
   +-----------+            +-----------+
        |                         |
        |                         |
   +----v------+            +-----v-----+
   |           |            |           |
   |    o      |            |     a     |
   |           |            |           |
   +-----------+            +-----------+
                                  |
                                  |
                            +-----v-----+
                            |           |
                            |     t     |
                            |           |
                            +-----------+
```

So in this example, we can see the `trie` with keys, `go` and `cat` . The compressed trie or `radix tree` differs from `trie` in that all intermediates nodes which have only one child are removed.

Radix tree in linux kernel is the data structure which maps values to integer keys. It is represented by the following structures from the file include/linux/radix-tree.h:

```
struct radix_tree_root {
         unsigned int            height;
         gfp_t                   gfp_mask;
         struct radix_tree_node  __rcu *rnode;
};
```

This structure presents the root of a radix tree and contains three fields:

* `height` - height of the tree;
* `gfp_mask` - tells how memory allocations will be performed;
* `rnode` - pointer to the child node.

The first field we will discuss is `gfp_mask` :

Low-level kernel memory allocation functions take a set of flags as - `gfp_mask` , which describes how that allocation is to be performed. These `GFP_` flags which control the allocation process can have following values: (`GF_NOIO` flag) means sleep and wait for memory, (`__GFP_HIGHMEM` flag) means high memory can be used, (`GFP_ATOMIC` flag) means the allocation process has high-priority and can't sleep etc.

* `GFP_NOIO` - can sleep and wait for memory;
* `__GFP_HIGHMEM` - high memory can be used;
* `GFP_ATOMIC` - allocation process is high-priority and can't sleep;

etc.

The next field is `rnode` :

```
struct radix_tree_node {
        unsigned int    path;
        unsigned int    count;
        union {
                struct {
                        struct radix_tree_node *parent;
                        void *private_data;
                };
                struct rcu_head rcu_head;
        };
        /* For tree user */
        struct list_head private_list;
        void __rcu      *slots[RADIX_TREE_MAP_SIZE];
        unsigned long   tags[RADIX_TREE_MAX_TAGS][RADIX_TREE_TAG_LONGS];
};
```

This structure contains information about the offset in a parent and height from the bottom, count of the child nodes and fields for accessing and freeing a node. These fields are described below:

* `path` - offset in parent & height from the bottom;
* `count` - count of the child nodes;
* `parent` - pointer to the parent node;
* `private_data` - used by the user of a tree;
* `rcu_head` - used for freeing a node;
* `private_list` - used by the user of a tree;

The two last fields of the `radix_tree_node` - `tags` and `slots` are important and interesting. Every node can contains a set of slots which store pointers to the data. Empty slots in the linux kernel radix tree implementation store `NULL` . Radix trees in the linux kernel also supports tags which are associated with the `tags` fields in the `radix_tree_node` structure. Tags allow individual bits to be set on records which are stored in the radix tree.

Now that we know about radix tree structure, it is time to look on its API.

Linux kernel radix tree API

We start from the data structure initialization. There are two ways to initialize a new radix tree. The first is to use `RADIX_TREE` macro:

```
RADIX_TREE(name, gfp_mask);
```

```
`
```

As you can see we pass the `name` parameter, so with the `RADIX_TREE` macro we can define and initialize radix tree with the given name. Implementation of the `RADIX_TREE` is easy:

```
#define RADIX_TREE(name, mask) \
         struct radix_tree_root name = RADIX_TREE_INIT(mask)

#define RADIX_TREE_INIT(mask)   { \
        .height = 0,              \
        .gfp_mask = (mask),       \
        .rnode = NULL,            \
}
```

At the beginning of the `RADIX_TREE` macro we define instance of the `radix_tree_root` structure with the given name and call `RADIX_TREE_INIT` macro with the given mask. The `RADIX_TREE_INIT` macro just initializes `radix_tree_root` structure with the default values and the given mask.

The second way is to define `radix_tree_root` structure by hand and pass it with mask to the `INIT_RADIX_TREE` macro:

```
struct radix_tree_root my_radix_tree;
INIT_RADIX_TREE(my_tree, gfp_mask_for_my_radix_tree);
```

where:

```
#define INIT_RADIX_TREE(root, mask)  \
do {                                 \
        (root)->height = 0;          \
        (root)->gfp_mask = (mask);   \
        (root)->rnode = NULL;        \
} while (0)
```

makes the same initialization with default values as it does `RADIX_TREE_INIT` macro.

The next are two functions for inserting and deleting records to/from a radix tree:

- `radix_tree_insert` ;
- `radix_tree_delete` ;

The first `radix_tree_insert` function takes three parameters:

- root of a radix tree;
- index key;
- data to insert;

The `radix_tree_delete` function takes the same set of parameters as the `radix_tree_insert` , but without data.

The search in a radix tree implemented in two ways:

- `radix_tree_lookup` ;
- `radix_tree_gang_lookup` ;
- `radix_tree_lookup_slot` .

The first `radix_tree_lookup` function takes two parameters:

- root of a radix tree;
- index key;

This function tries to find the given key in the tree and return the record associated with this key. The second `radix_tree_gang_lookup` function have the following signature

```
unsigned int radix_tree_gang_lookup(struct radix_tree_root *root,
                                    void **results,
                                    unsigned long first_index,
                                    unsigned int max_items);
```

and returns number of records, sorted by the keys, starting from the first index. Number of the returned records will not be greater than `max_items` value.

And the last `radix_tree_lookup_slot` function will return the slot which will contain the data.

Links

- Radix tree
- Trie

Data Structures in the Linux Kernel

Bit arrays and bit operations in the Linux kernel

Besides different linked and tree based data structures, the Linux kernel provides API for bit arrays or `bitmap` . Bit arrays are heavily used in the Linux kernel and following source code files contain common `API` for work with such structures:

- lib/bitmap.c
- include/linux/bitmap.h

Besides these two files, there is also architecture-specific header file which provides optimized bit operations for certain architecture. We consider x86_64 architecture, so in our case it will be:

- arch/x86/include/asm/bitops.h

header file. As I just wrote above, the `bitmap` is heavily used in the Linux kernel. For example a `bit array` is used to store set of online/offline processors for systems which support hot-plug cpu (more about this you can read in the cpumasks part), a `bit array` stores set of allocated irqs during initialization of the Linux kernel and etc.

So, the main goal of this part is to see how `bit arrays` are implemented in the Linux kernel. Let's start.

Declaration of bit array

Before we will look on `API` for bitmaps manipulation, we must know how to declare it in the Linux kernel. There are two common method to declare own bit array. The first simple way to declare a bit array is to array of `unsigned long` . For example:

```
unsigned long my_bitmap[8]
```

The second way is to use the `DECLARE_BITMAP` macro which is defined in the include/linux/types.h header file:

```
#define DECLARE_BITMAP(name,bits) \
    unsigned long name[BITS_TO_LONGS(bits)]
```

We can see that `DECLARE_BITMAP` macro takes two parameters:

- `name` - name of bitmap;
- `bits` - amount of bits in bitmap;

and just expands to the definition of `unsigned long` array with `BITS_TO_LONGS(bits)` elements, where the `BITS_TO_LONGS` macro converts a given number of bits to number of `longs` or in other words it calculates how many `8` byte elements in `bits` :

```
#define BITS_PER_BYTE           8
#define DIV_ROUND_UP(n,d) (((n) + (d) - 1) / (d))
#define BITS_TO_LONGS(nr)       DIV_ROUND_UP(nr, BITS_PER_BYTE * sizeof(long))
```

So, for example `DECLARE_BITMAP(my_bitmap, 64)` will produce:

```
>>> (((64) + (64) - 1) / (64))
1
```

and:

```
unsigned long my_bitmap[1];
```

After we are able to declare a bit array, we can start to use it.

Architecture-specific bit operations

We already saw above a couple of source code and header files which provide API for manipulation of bit arrays. The most important and widely used API of bit arrays is architecture-specific and located as we already know in the arch/x86/include/asm/bitops.h header file.

First of all let's look at the two most important functions:

- `set_bit` ;
- `clear_bit` .

I think that there is no need to explain what these function do. This is already must be clear from their name. Let's look on their implementation. If you will look into the arch/x86/include/asm/bitops.h header file, you will note that each of these functions represented by two variants: atomic and not. Before we will start to dive into implementations of these functions, first of all we must to know a little about `atomic` operations.

In simple words atomic operations guarantees that two or more operations will not be performed on the same data concurrently. The `x86` architecture provides a set of atomic instructions, for example xchg instruction, cmpxchg instruction and etc. Besides atomic instructions, some of non-atomic instructions can be made atomic with the help of the lock instruction. It is enough to know about atomic operations for now, so we can begin to consider implementation of `set_bit` and `clear_bit` functions.

First of all, let's start to consider `non-atomic` variants of this function. Names of non-atomic `set_bit` and `clear_bit` starts from double underscore. As we already know, all of these functions are defined in the arch/x86/include/asm/bitops.h header file and the first function is `__set_bit` :

```
static inline void __set_bit(long nr, volatile unsigned long *addr)
{
    asm volatile("bts %1,%0" : ADDR : "Ir" (nr) : "memory");
}
```

As we can see it takes two arguments:

- `nr` - number of bit in a bit array.
- `addr` - address of a bit array where we need to set bit.

Note that the `addr` parameter is defined with `volatile` keyword which tells to compiler that value maybe changed by the given address. The implementation of the `__set_bit` is pretty easy. As we can see, it just contains one line of inline assembler code. In our case we are using the bts instruction which selects a bit which is specified with the first operand (`nr` in our case) from the bit array, stores the value of the selected bit in the CF flags register and set this bit.

Note that we can see usage of the `nr` , but there is `addr` here. You already might guess that the secret is in `ADDR` . The `ADDR` is the macro which is defined in the same header code file and expands to the string which contains value of the given address and `+m` constraint:

```
#define ADDR                BITOP_ADDR(addr)
#define BITOP_ADDR(x) "+m" (*(volatile long *) (x))
```

Besides the `+m` , we can see other constraints in the `__set_bit` function. Let's look on they and try to understand what do they mean:

- `+m` - represents memory operand where `+` tells that the given operand will be input and output operand;
- `I` - represents integer constant;
- `r` - represents register operand

Besides these constraint, we also can see - the `memory` keyword which tells compiler that this code will change value in memory. That's all. Now let's look at the same function but at `atomic` variant. It looks more complex that its `non-atomic` variant:

```
static __always_inline void
set_bit(long nr, volatile unsigned long *addr)
{
    if (IS_IMMEDIATE(nr)) {
        asm volatile(LOCK_PREFIX "orb %1,%0"
            : CONST_MASK_ADDR(nr, addr)
            : "iq" ((u8)CONST_MASK(nr))
            : "memory");
    } else {
        asm volatile(LOCK_PREFIX "bts %1,%0"
            : BITOP_ADDR(addr) : "Ir" (nr) : "memory");
    }
}
```

First of all note that this function takes the same set of parameters that `__set_bit` , but additionally marked with the `__always_inline` attribute. The `__always_inline` is macro which defined in the include/linux/compiler-gcc.h and just expands to the `always_inline` attribute:

```
#define __always_inline inline __attribute__((always_inline))
```

which means that this function will be always inlined to reduce size of the Linux kernel image. Now let's try to understand implementation of the `set_bit` function. First of all we check a given number of bit at the beginning of the `set_bit` function. The `IS_IMMEDIATE` macro defined in the same header file and expands to the call of the builtin gcc function:

```
#define IS_IMMEDIATE(nr)        (__builtin_constant_p(nr))
```

The `__builtin_constant_p` builtin function returns `1` if the given parameter is known to be constant at compile-time and returns `0` in other case. We no need to use slow `bts` instruction to set bit if the given number of bit is known in compile time constant. We can just apply bitwise or for byte from the give address which contains given bit and masked number of bits where high bit is `1` and other is zero. In other case if the given number of bit is not known constant at compile-time, we do the same as we did in the `__set_bit` function. The `CONST_MASK_ADDR` macro:

```
#define CONST_MASK_ADDR(nr, addr)    BITOP_ADDR((void *)(addr) + ((nr)>>3))
```

expands to the give address with offset to the byte which contains a given bit. For example we have address `0x1000` and the number of bit is `0x9` . So, as `0x9` is `one byte + one bit` our address with be `addr + 1` :

```
>>> hex(0x1000 + (0x9 >> 3))
'0x1001'
```

The `CONST_MASK` macro represents our given number of bit as byte where high bit is `1` and other bits are `0` :

```
#define CONST_MASK(nr)            (1 << ((nr) & 7))
```

```
>>> bin(1 << (0x9 & 7))
'0b10'
```

In the end we just apply bitwise `or` for these values. So, for example if our address will be `0x4097` and we need to set `0x9` bit:

```
>>> bin(0x4097)
'0b100000010010111'
>>> bin((0x4097 >> 0x9) | (1 << (0x9 & 7)))
'0b100010'
```

the `ninth` bit will be set.

Note that all of these operations are marked with `LOCK_PREFIX` which is expands to the lock instruction which guarantees atomicity of this operation.

As we already know, besides the `set_bit` and `__set_bit` operations, the Linux kernel provides two inverse functions to clear bit in atomic and non-atomic context. They are `clear_bit` and `__clear_bit`. Both of these functions are defined in the same header file and takes the same set of arguments. But not only arguments are similar. Generally these functions are very similar on the `set_bit` and `__set_bit`. Let's look on the implementation of the non-atomic `__clear_bit` function:

```
static inline void __clear_bit(long nr, volatile unsigned long *addr)
{
    asm volatile("btr %1,%0" : ADDR : "Ir" (nr));
}
```

Yes. As we see, it takes the same set of arguments and contains very similar block of inline assembler. It just uses the btr instruction instead of `bts`. As we can understand form the function's name, it clears a given bit by the given address. The `btr` instruction acts like `bts`. This instruction also selects a given bit which is specified in the first operand, stores its value in the `CF` flag register and clears this bit in the given bit array which is specified with second operand.

The atomic variant of the `__clear_bit` is `clear_bit`:

```
static __always_inline void
clear_bit(long nr, volatile unsigned long *addr)
{
    if (IS_IMMEDIATE(nr)) {
        asm volatile(LOCK_PREFIX "andb %1,%0"
            : CONST_MASK_ADDR(nr, addr)
            : "iq" ((u8)~CONST_MASK(nr)));
    } else {
        asm volatile(LOCK_PREFIX "btr %1,%0"
            : BITOP_ADDR(addr)
            : "Ir" (nr));
    }
}
```

and as we can see it is very similar on `set_bit` and just contains two differences. The first difference it uses `btr` instruction to clear bit when the `set_bit` uses `bts` instruction to set bit. The second difference it uses negated mask and `and` instruction to clear bit in the given byte when the `set_bit` uses `or` instruction.

That's all. Now we can set and clear bit in any bit array and and we can go to other operations on bitmasks.

Most widely used operations on a bit arrays are set and clear bit in a bit array in the Linux kernel. But besides this operations it is useful to do additional operations on a bit array. Yet another widely used operation in the Linux kernel - is to know is a given bit set or not in a bit array. We can achieve this with the help of the `test_bit` macro. This macro is defined in the arch/x86/include/asm/bitops.h header file and expands to the call of the `constant_test_bit` or `variable_test_bit` depends on bit number:

```
#define test_bit(nr, addr)            \
    (__builtin_constant_p((nr))                  \
     ? constant_test_bit((nr), (addr))           \
     : variable_test_bit((nr), (addr)))
```

So, if the `nr` is known in compile time constant, the `test_bit` will be expanded to the call of the `constant_test_bit` function or `variable_test_bit` in other case. Now let's look at implementations of these functions. Let's start from the `variable_test_bit` :

```
static inline int variable_test_bit(long nr, volatile const unsigned long *addr)
{
    int oldbit;

    asm volatile("bt %2,%1\n\t"
             "sbb %0,%0"
             : "=r" (oldbit)
             : "m" (*(unsigned long *)addr), "Ir" (nr));

    return oldbit;
}
```

The `variable_test_bit` function takes similar set of arguments as `set_bit` and other function take. We also may see inline assembly code here which executes bt and sbb instruction. The `bt` or `bit test` instruction selects a given bit which is specified with first operand from the bit array which is specified with the second operand and stores its value in the CF bit of flags register. The second `sbb` instruction subtracts first operand from second and subtracts value of the `CF` . So, here write a value of a given bit number from a given bit array to the `CF` bit of flags register and execute `sbb` instruction which calculates: `00000000 - CF` and writes the result to the `oldbit` .

The `constant_test_bit` function does the same as we saw in the `set_bit` :

```
static __always_inline int constant_test_bit(long nr, const volatile unsigned long *addr)
{
    return ((1UL << (nr & (BITS_PER_LONG-1))) &
        (addr[nr >> _BITOPS_LONG_SHIFT])) != 0;
}
```

It generates a byte where high bit is `1` and other bits are `0` (as we saw in `CONST_MASK`) and applies bitwise and to the byte which contains a given bit number.

The next widely used bit array related operation is to change bit in a bit array. The Linux kernel provides two helper for this:

- `__change_bit` ;
- `change_bit` .

As you already can guess, these two variants are atomic and non-atomic as for example `set_bit` and `__set_bit` . For the start, let's look at the implementation of the `__change_bit` function:

```
static inline void __change_bit(long nr, volatile unsigned long *addr)
{
    asm volatile("btc %1,%0" : ADDR : "Ir" (nr));
}
```

Pretty easy, is not it? The implementation of the `__change_bit` is the same as `__set_bit` , but instead of `bts` instruction, we are using btc. This instruction selects a given bit from a given bit array, stores its value in the `CF` and changes its value by the applying of complement operation. So, a bit with value `1` will be `0` and vice versa:

```
>>> int(not 1)
0
>>> int(not 0)
1
```

The atomic version of the `__change_bit` is the `change_bit` function:

```
static inline void change_bit(long nr, volatile unsigned long *addr)
{
    if (IS_IMMEDIATE(nr)) {
        asm volatile(LOCK_PREFIX "xorb %1,%0"
            : CONST_MASK_ADDR(nr, addr)
            : "iq" ((u8)CONST_MASK(nr)));
    } else {
        asm volatile(LOCK_PREFIX "btc %1,%0"
            : BITOP_ADDR(addr)
            : "Ir" (nr));
    }
}
```

It is similar on `set_bit` function, but also has two differences. The first difference is `xor` operation instead of `or` and the second is `btc` instead of `bts` .

For this moment we know the most important architecture-specific operations with bit arrays. Time to look at generic bitmap API.

Common bit operations

Besides the architecture-specific API from the arch/x86/include/asm/bitops.h header file, the Linux kernel provides common API for manipulation of bit arrays. As we know from the beginning of this part, we can find it in the include/linux/bitmap.h header file and additionally in the * lib/bitmap.c source code file. But before these source code files let's look into the include/linux/bitops.h header file which provides a set of useful macro. Let's look on some of they.

First of all let's look at following four macros:

- `for_each_set_bit`
- `for_each_set_bit_from`
- `for_each_clear_bit`
- `for_each_clear_bit_from`

All of these macros provide iterator over certain set of bits in a bit array. The first macro iterates over bits which are set, the second does the same, but starts from a certain bits. The last two macros do the same, but iterates over clear bits. Let's look on implementation of the `for_each_set_bit` macro:

```
#define for_each_set_bit(bit, addr, size) \
    for ((bit) = find_first_bit((addr), (size));        \
         (bit) < (size);                    \
         (bit) = find_next_bit((addr), (size), (bit) + 1))
```

As we may see it takes three arguments and expands to the loop from first set bit which is returned as result of the `find_first_bit` function and to the last bit number while it is less than given size.

Besides these four macros, the arch/x86/include/asm/bitops.h provides API for rotation of `64-bit` or `32-bit` values and etc.

The next header file which provides API for manipulation with a bit arrays. For example it provides two functions:

- `bitmap_zero` ;
- `bitmap_fill` .

To clear a bit array and fill it with `1`. Let's look on the implementation of the `bitmap_zero` function:

```
static inline void bitmap_zero(unsigned long *dst, unsigned int nbits)
{
    if (small_const_nbits(nbits))
        *dst = 0UL;
    else {
        unsigned int len = BITS_TO_LONGS(nbits) * sizeof(unsigned long);
        memset(dst, 0, len);
    }
}
```

First of all we can see the check for `nbits`. The `small_const_nbits` is macro which defined in the same header file and looks:

```
#define small_const_nbits(nbits) \
    (__builtin_constant_p(nbits) && (nbits) <= BITS_PER_LONG)
```

As we may see it checks that `nbits` is known constant in compile time and `nbits` value does not overflow `BITS_PER_LONG` or `64`. If bits number does not overflow amount of bits in a `long` value we can just set to zero. In other case we need to calculate how many `long` values do we need to fill our bit array and fill it with memset.

The implementation of the `bitmap_fill` function is similar on implementation of the `biramp_zero` function, except we fill a given bit array with `0xff` values or `0b11111111`:

```
static inline void bitmap_fill(unsigned long *dst, unsigned int nbits)
{
    unsigned int nlongs = BITS_TO_LONGS(nbits);
    if (!small_const_nbits(nbits)) {
        unsigned int len = (nlongs - 1) * sizeof(unsigned long);
        memset(dst, 0xff,  len);
    }
    dst[nlongs - 1] = BITMAP_LAST_WORD_MASK(nbits);
}
```

Besides the `bitmap_fill` and `bitmap_zero` functions, the include/linux/bitmap.h header file provides `bitmap_copy` which is similar on the `bitmap_zero`, but just uses memcpy instead of memset. Also it provides bitwise operations for bit array like `bitmap_and`, `bitmap_or`, `bitamp_xor` and etc. We will not consider implementation of these functions because it is easy to understand implementations of these functions if you understood all from this part. Anyway if you are interested how did these function implemented, you may open include/linux/bitmap.h header file and start to research.

That's all.

Links

- bitmap
- linked data structures
- tree data structures
- hot-plug
- cpumasks
- IRQs
- API
- atomic operations
- xchg instruction
- cmpxchg instruction
- lock instruction

- bts instruction
- btr instruction
- bt instruction
- sbb instruction
- btc instruction
- man memcpy
- man memset
- CF
- inline assembler
- gcc

Theory

This chapter describes various theoretical concepts and concepts which are not directly related to practice but useful to know.

- Paging
- Elf64 format
- Inline assembly

Paging

Introduction

In the fifth part of the series `Linux kernel booting process` we learned about what the kernel does in its earliest stage. In the next step the kernel will initialize different things like `initrd` mounting, lockdep initialization, and many many other things, before we can see how the kernel runs the first init process.

Yeah, there will be many different things, but many many and once again many work with **memory**.

In my view, memory management is one of the most complex parts of the Linux kernel and in system programming in general. This is why we need to get acquainted with paging, before we proceed with the kernel initialization stuff.

`Paging` is a mechanism that translates a linear memory address to a physical address. If you have read the previous parts of this book, you may remember that we saw segmentation in real mode when physical addresses are calculated by shifting a segment register by four and adding an offset. We also saw segmentation in protected mode, where we used the descriptor tables and base addresses from descriptors with offsets to calculate the physical addresses. Now we will see paging in 64-bit mode.

As the Intel manual says:

> Paging provides a mechanism for implementing a conventional demand-paged, virtual-memory system where sections of a program's execution environment are mapped into physical memory as needed.

So... In this post I will try to explain the theory behind paging. Of course it will be closely related to the `x86_64` version of the Linux kernel, but we will not go into too much details (at least in this post).

Enabling paging

There are three paging modes:

- 32-bit paging;
- PAE paging;
- IA-32e paging.

We will only explain the last mode here. To enable the `IA-32e paging` paging mode we need to do following things:

- set the `CR0.PG` bit;
- set the `CR4.PAE` bit;
- set the `IA32_EFER.LME` bit.

We already saw where those bits were set in arch/x86/boot/compressed/head_64.S:

```
movl    $(X86_CR0_PG | X86_CR0_PE), %eax
movl    %eax, %cr0
```

and

```
movl    $MSR_EFER, %ecx
rdmsr
btsl    $_EFER_LME, %eax
wrmsr
```

Paging structures

Paging structures

Paging divides the linear address space into fixed-size pages. Pages can be mapped into the physical address space or external storage. This fixed size is `4096` bytes for the `x86_64` Linux kernel. To perform the translation from linear address to physical address, special structures are used. Every structure is `4096` bytes and contains `512` entries (this only for `PAE` and `IA32_EFER.LME` modes). Paging structures are hierarchical and the Linux kernel uses 4 level of paging in the `x86_64` architecture. The CPU uses a part of linear addresses to identify the entry in another paging structure which is at the lower level, physical memory region (`page frame`) or physical address in this region (`page offset`). The address of the top level paging structure located in the `cr3` register. We have already seen this in arch/x86/boot/compressed/head_64.S:

```
leal    pgtable(%ebx), %eax
movl    %eax, %cr3
```

We build the page table structures and put the address of the top-level structure in the `cr3` register. Here `cr3` is used to store the address of the top-level structure, the `PML4` or `Page Global Directory` as it is called in the Linux kernel. `cr3` is 64-bit register and has the following structure:

```
63                  52 51                                                    32
 --------------------------------------------------------------------------------
|                     |                                                          |
|    Reserved MBZ     |            Address of the top level structure            |
|                     |                                                          |
 --------------------------------------------------------------------------------
31                                  12 11            5     4     3 2             0
 --------------------------------------------------------------------------------
|                                     |               |  P  |  P  |              |
|  Address of the top level structure |    Reserved   |  C  |  W  |    Reserved  |
|                                     |               |  D  |  T  |              |
 --------------------------------------------------------------------------------
```

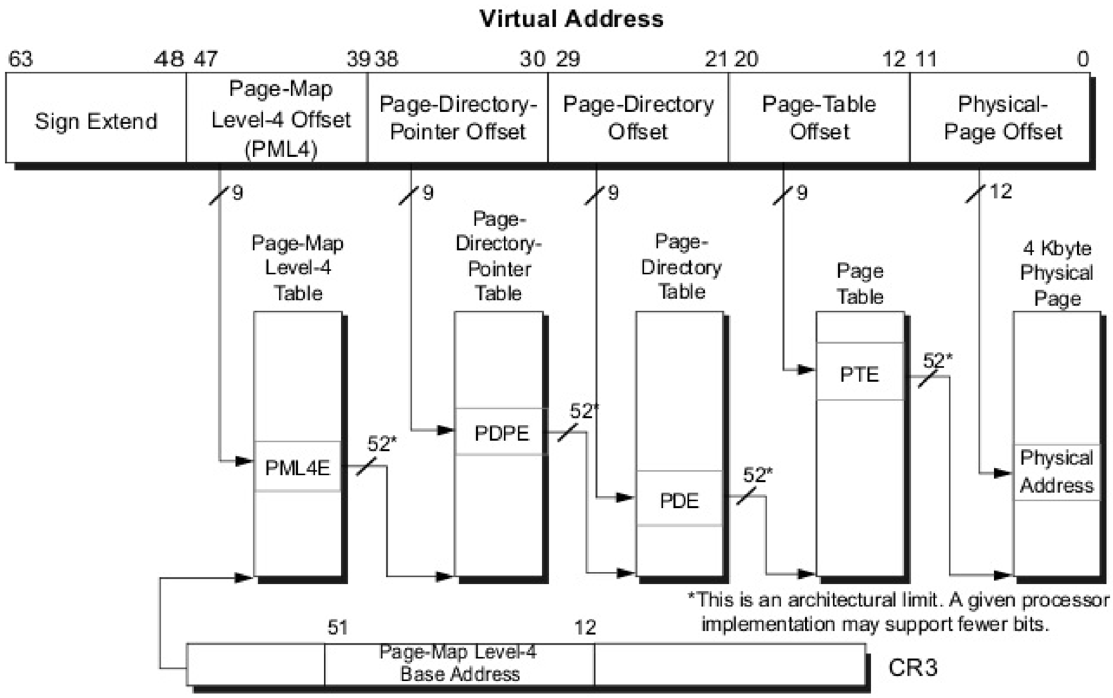
These fields have the following meanings:

- Bits 63:52 - reserved must be 0.
- Bits 51:12 - stores the address of the top level paging structure;
- Reserved - reserved must be 0;
- Bits 4 : 3 - PWT or Page-Level Writethrough and PCD or Page-level cache disable indicate. These bits control the way the page or Page Table is handled by the hardware cache;
- Bits 2 : 0 - ignored;

The linear address translation is following:

- A given linear address arrives to the MMU instead of memory bus.
- 64-bit linear address is split into some parts. Only low 48 bits are significant, it means that `2^48` or 256 TBytes of linear-address space may be accessed at any given time.
- `cr3` register stores the address of the 4 top-level paging structure.
- `47:39` bits of the given linear address store an index into the paging structure level-4, `38:30` bits store index into the paging structure level-3, `29:21` bits store an index into the paging structure level-2, `20:12` bits store an index into the paging structure level-1 and `11:0` bits provide the offset into the physical page in byte.

schematically, we can imagine it like this:



Every access to a linear address is either a supervisor-mode access or a user-mode access. This access is determined by the `CPL` (current privilege level). If `CPL < 3` it is a supervisor mode access level, otherwise it is a user mode access level. For example, the top level page table entry contains access bits and has the following structure:

```
63  62                  52 51                                                    32
 --------------------------------------------------------------------------------
| N |                     |                                                     |
|   |     Available       |     Address of the paging structure on lower level  |
| X |                     |                                                     |
 --------------------------------------------------------------------------------
31                                              12 11   9 8 7 6 5   4   3 2 1     0
 --------------------------------------------------------------------------------
|                                                |     | M |I| | P | P |U|W|     |
| Address of the paging structure on lower level | AVL | B |G|A| C | W | | |  P |
|                                                |     | Z |N| | D | T |S|R|     |
 --------------------------------------------------------------------------------
```

Where:

- 63 bit - N/X bit (No Execute Bit) which presents ability to execute the code from physical pages mapped by the table entry;
- 62:52 bits - ignored by CPU, used by system software;
- 51:12 bits - stores physical address of the lower level paging structure;
- 11: 9 bits - ignored by CPU;
- MBZ - must be zero bits;
- Ignored bits;
- A - accessed bit indicates was physical page or page structure accessed;
- PWT and PCD used for cache;
- U/S - user/supervisor bit controls user access to all the physical pages mapped by this table entry;
- R/W - read/write bit controls read/write access to all the physical pages mapped by this table entry;
- P - present bit. Current bit indicates was page table or physical page loaded into primary memory or not.

Ok, we know about the paging structures and their entries. Now let's see some details about 4-level paging in the Linux kernel.

Paging structures in the Linux kernel

As we've seen, the Linux kernel in `x86_64` uses 4-level page tables. Their names are:

- Page Global Directory
- Page Upper Directory
- Page Middle Directory
- Page Table Entry

After you've compiled and installed the Linux kernel, you can see the `System.map` file which stores the virtual addresses of the functions that are used by the kernel. For example:

```
$ grep "start_kernel" System.map
ffffffff81efe497 T x86_64_start_kernel
ffffffff81efeaa2 T start_kernel
```

We can see `0xffffffff81efe497` here. I doubt you really have that much RAM installed. But anyway, `start_kernel` and `x86_64_start_kernel` will be executed. The address space in `x86_64` is `2^64` wide, but it's too large, that's why a smaller address space is used, only 48-bits wide. So we have a situation where the physical address space is limited to 48 bits, but addressing still performs with 64 bit pointers. How is this problem solved? Look at this diagram:

```
0xffffffffffffffff  +-----------+
                    |           |
                    |           | Kernelspace
                    |           |
0xffff800000000000  +-----------+
                    |           |
                    |           |
                    |   hole    |
                    |           |
                    |           |
0x00007fffffffffff  +-----------+
                    |           |
                    |           |  Userspace
                    |           |
0x0000000000000000  +-----------+
```

This solution is `sign extension`. Here we can see that the lower 48 bits of a virtual address can be used for addressing. Bits `63:48` can be either only zeroes or only ones. Note that the virtual address space is split into 2 parts:

- Kernel space
- Userspace

Userspace occupies the lower part of the virtual address space, from `0x000000000000000` to `0x00007fffffffffff` and kernel space occupies the highest part from `0xffff8000000000` to `0xffffffffffffffff`. Note that bits `63:48` is 0 for userspace and 1 for kernel space. All addresses which are in kernel space and in userspace or in other words which higher `63:48` bits are zeroes or ones are called `canonical` addresses. There is a `non-canonical` area between these memory regions. Together these two memory regions (kernel space and user space) are exactly `2^48` bits wide. We can find the virtual memory map with 4 level page tables in the Documentation/x86/x86_64/mm.txt:

```
0000000000000000 - 00007fffffffffff (=47 bits) user space, different per mm
hole caused by [48:63] sign extension
ffff800000000000 - ffff87ffffffffff (=43 bits) guard hole, reserved for hypervisor
ffff880000000000 - ffffc7ffffffffff (=64 TB) direct mapping of all phys. memory
ffffc80000000000 - ffffc8ffffffffff (=40 bits) hole
ffffc90000000000 - ffffe8ffffffffff (=45 bits) vmalloc/ioremap space
ffffe90000000000 - ffffe9ffffffffff (=40 bits) hole
ffffea0000000000 - ffffeaffffffffff (=40 bits) virtual memory map (1TB)
... unused hole ...
```

```
ffffec0000000000 - fffffc0000000000 (=44 bits) kasan shadow memory (16TB)
... unused hole ...
ffffff0000000000 - ffffff7fffffffff (=39 bits) %esp fixup stacks
... unused hole ...
ffffffff80000000 - ffffffffa0000000 (=512 MB)  kernel text mapping, from phys 0
ffffffffa0000000 - ffffffffff5fffff (=1525 MB) module mapping space
ffffffffff600000 - ffffffffffdfffff (=8 MB) vsyscalls
ffffffffffe00000 - ffffffffffffffff (=2 MB) unused hole
```

We can see here the memory map for user space, kernel space and the non-canonical area in-between them. The user space memory map is simple. Let's take a closer look at the kernel space. We can see that it starts from the guard hole which is reserved for the hypervisor. We can find the definition of this guard hole in arch/x86/include/asm/page_64_types.h:

```
#define __PAGE_OFFSET _AC(0xffff880000000000, UL)
```

Previously this guard hole and `__PAGE_OFFSET` was from `0xffff800000000000` to `0xffff87ffffffffff` to prevent access to non-canonical area, but was later extended by 3 bits for the hypervisor.

Next is the lowest usable address in kernel space - `ffff880000000000`. This virtual memory region is for direct mapping of all the physical memory. After the memory space which maps all the physical addresses, the guard hole. It needs to be between the direct mapping of all the physical memory and the vmalloc area. After the virtual memory map for the first terabyte and the unused hole after it, we can see the `kasan` shadow memory. It was added by commit and provides the kernel address sanitizer. After the next unused hole we can see the `esp` fixup stacks (we will talk about it in other parts of this book) and the start of the kernel text mapping from the physical address - `0`. We can find the definition of this address in the same file as the `__PAGE_OFFSET`:

```
#define __START_KERNEL_map      _AC(0xffffffff80000000, UL)
```

Usually kernel's `.text` starts here with the `CONFIG_PHYSICAL_START` offset. We have seen it in the post about ELF64:

```
readelf -s vmlinux | grep ffffffff81000000
     1: ffffffff81000000     0 SECTION LOCAL  DEFAULT    1
 65099: ffffffff81000000     0 NOTYPE  GLOBAL DEFAULT    1 _text
 90766: ffffffff81000000     0 NOTYPE  GLOBAL DEFAULT    1 startup_64
```

Here I check `vmlinux` with `CONFIG_PHYSICAL_START` is `0x1000000`. So we have the start point of the kernel `.text` - `0xffffffff80000000` and offset - `0x1000000`, the resulted virtual address will be `0xffffffff80000000 + 1000000 = 0xffffffff81000000`.

After the kernel `.text` region there is the virtual memory region for kernel module, `vsyscalls` and an unused hole of 2 megabytes.

We've seen how virtual memory map in the kernel is laid out and how a virtual address is translated into a physical one. Let's take the following address as example:

```
0xffffffff81000000
```

In binary it will be:

```
1111111111111111 111111111 111111110 000001000 000000000 000000000000
      63:48        47:39     38:30     29:21     20:12      11:0
```

This virtual address is split in parts as described above:

- `63:48` - bits not used;

* `47:39` - bits store an index into the paging structure level-4;
* `38:30` - bits store index into the paging structure level-3;
* `29:21` - bits store an index into the paging structure level-2;
* `20:12` - bits store an index into the paging structure level-1;
* `11:0` - bits provide the offset into the physical page in byte.

That is all. Now you know a little about theory of `paging` and we can go ahead in the kernel source code and see the first initialization steps.

Conclusion

It's the end of this short part about paging theory. Of course this post doesn't cover every detail of paging, but soon we'll see in practice how the Linux kernel builds paging structures and works with them.

Please note that English is not my first language and I am really sorry for any inconvenience. If you've found any mistakes please send me PR to linux-insides.

Links

* Paging on Wikipedia
* Intel 64 and IA-32 architectures software developer's manual volume 3A
* MMU
* ELF64
* Documentation/x86/x86_64/mm.txt
* Last part - Kernel booting process

Executable and Linkable Format

ELF (Executable and Linkable Format) is a standard file format for executable files, object code, shared libraries and core dumps. Linux and many UNIX-like operating systems use this format. Let's look at the structure of the ELF-64 Object File Format and some definitions in the linux kernel source code which related with it.

An ELF object file consists of the following parts:

- ELF header - describes the main characteristics of the object file: type, CPU architecture, the virtual address of the entry point, the size and offset of the remaining parts, etc...;
- Program header table - lists the available segments and their attributes. Program header table need loaders for placing sections of the file as virtual memory segments;
- Section header table - contains the description of the sections.

Now let's have a closer look on these components.

ELF header

The ELF header is located at the beginning of the object file. Its main purpose is to locate all other parts of the object file. The File header contains the following fields:

- ELF identification - array of bytes which helps identify the file as an ELF object file and also provides information about general object file characteristic;
- Object file type - identifies the object file type. This field can describe that ELF file is a relocatable object file, an executable file, etc...;
- Target architecture;
- Version of the object file format;
- Virtual address of the program entry point;
- File offset of the program header table;
- File offset of the section header table;
- Size of an ELF header;
- Size of a program header table entry;
- and other fields...

You can find the `elf64_hdr` structure which presents ELF64 header in the linux kernel source code:

```
typedef struct elf64_hdr {
    unsigned char    e_ident[EI_NIDENT];
    Elf64_Half e_type;
    Elf64_Half e_machine;
    Elf64_Word e_version;
    Elf64_Addr e_entry;
    Elf64_Off e_phoff;
    Elf64_Off e_shoff;
    Elf64_Word e_flags;
    Elf64_Half e_ehsize;
    Elf64_Half e_phentsize;
    Elf64_Half e_phnum;
    Elf64_Half e_shentsize;
    Elf64_Half e_shnum;
    Elf64_Half e_shstrndx;
} Elf64_Ehdr;
```

This structure defined in the elf.h

Sections

All data stores in a sections in an Elf object file. Sections identified by index in the section header table. Section header contains

All data stores in a sections in an Elf object file. Sections identified by index in the section header table. Section header contains following fields:

- Section name;
- Section type;
- Section attributes;
- Virtual address in memory;
- Offset in file;
- Size of section;
- Link to other section;
- Miscellaneous information;
- Address alignment boundary;
- Size of entries, if section has table;

And presented with the following `elf64_shdr` structure in the linux kernel:

```
typedef struct elf64_shdr {
    Elf64_Word sh_name;
    Elf64_Word sh_type;
    Elf64_Xword sh_flags;
    Elf64_Addr sh_addr;
    Elf64_Off sh_offset;
    Elf64_Xword sh_size;
    Elf64_Word sh_link;
    Elf64_Word sh_info;
    Elf64_Xword sh_addralign;
    Elf64_Xword sh_entsize;
} Elf64_Shdr;
```

elf.h

Program header table

All sections are grouped into segments in an executable or shared object file. Program header is an array of structures which describe every segment. It looks like:

```
typedef struct elf64_phdr {
    Elf64_Word p_type;
    Elf64_Word p_flags;
    Elf64_Off p_offset;
    Elf64_Addr p_vaddr;
    Elf64_Addr p_paddr;
    Elf64_Xword p_filesz;
    Elf64_Xword p_memsz;
    Elf64_Xword p_align;
} Elf64_Phdr;
```

in the linux kernel source code.

`elf64_phdr` defined in the same elf.h.

The ELF object file also contains other fields/structures which you can find in the Documentation. Now let's a look at the `vmlinux` ELF object.

vmlinux

`vmlinux` is also a relocatable ELF object file . We can take a look at it with the `readelf` util. First of all let's look at the header:

```
$ readelf -h  vmlinux
ELF Header:
  Magic:   7f 45 4c 46 02 01 01 00 00 00 00 00 00 00 00 00
  Class:                             ELF64
  Data:                              2's complement, little endian
  Version:                           1 (current)
  OS/ABI:                            UNIX - System V
  ABI Version:                       0
  Type:                              EXEC (Executable file)
  Machine:                           Advanced Micro Devices X86-64
  Version:                           0x1
  Entry point address:               0x1000000
  Start of program headers:          64 (bytes into file)
  Start of section headers:          381608416 (bytes into file)
  Flags:                             0x0
  Size of this header:               64 (bytes)
  Size of program headers:           56 (bytes)
  Number of program headers:         5
  Size of section headers:           64 (bytes)
  Number of section headers:         73
  Section header string table index: 70
```

Here we can see that `vmlinux` is a 64-bit executable file.

We can read from the Documentation/x86/x86_64/mm.txt:

```
ffffffff80000000 - ffffffffa0000000 (=512 MB)  kernel text mapping, from phys 0
```

We can then look this address up in the `vmlinux` ELF object with:

```
$ readelf -s vmlinux | grep ffffffff81000000
     1: ffffffff81000000     0 SECTION LOCAL  DEFAULT    1
 65099: ffffffff81000000     0 NOTYPE  GLOBAL DEFAULT    1 _text
 90766: ffffffff81000000     0 NOTYPE  GLOBAL DEFAULT    1 startup_64
```

Note that the address of the `startup_64` routine is not `ffffffff80000000`, but `ffffffff81000000` and now I'll explain why.

We can see following definition in the arch/x86/kernel/vmlinux.lds.S:

```
    . = __START_KERNEL;
    ...
    ...
    ..
    /* Text and read-only data */
    .text :  AT(ADDR(.text) - LOAD_OFFSET) {
        _text = .;
        ...
        ...
        ...
    }
```

Where `__START_KERNEL` is:

```
#define __START_KERNEL        (__START_KERNEL_map + __PHYSICAL_START)
```

`__START_KERNEL_map` is the value from the documentation - `ffffffff80000000` and `__PHYSICAL_START` is `0x1000000`. That's why address of the `startup_64` is `ffffffff81000000`.

And at last we can get program headers from `vmlinux` with the following command:

```
readelf -l vmlinux

Elf file type is EXEC (Executable file)
Entry point 0x1000000
There are 5 program headers, starting at offset 64

Program Headers:
  Type           Offset             VirtAddr           PhysAddr
                 FileSiz            MemSiz              Flags  Align
  LOAD           0x0000000000200000 0xffffffff81000000 0x0000000001000000
                 0x0000000000cfd000 0x0000000000cfd000  R E    200000
  LOAD           0x0000000001000000 0xffffffff81e00000 0x0000000001e00000
                 0x0000000000100000 0x0000000000100000  RW     200000
  LOAD           0x0000000001200000 0x0000000000000000 0x0000000001f00000
                 0x0000000000014d98 0x0000000000014d98  RW     200000
  LOAD           0x0000000001315000 0xffffffff81f15000 0x0000000001f15000
                 0x000000000011d000 0x0000000000279000  RWE    200000
  NOTE           0x0000000000b17284 0xffffffff81917284 0x0000000001917284
                 0x0000000000000024 0x0000000000000024         4

 Section to Segment mapping:
  Segment Sections...
   00     .text .notes __ex_table .rodata __bug_table .pci_fixup .builtin_fw
          .tracedata __ksymtab __ksymtab_gpl __kcrctab __kcrctab_gpl
          __ksymtab_strings __param __modver
   01     .data .vvar
   02     .data..percpu
   03     .init.text .init.data .x86_cpu_dev.init .altinstructions
          .altinstr_replacement .iommu_table .apicdrivers .exit.text
          .smp_locks .data_nosave .bss .brk
```

Here we can see five segments with sections list. You can find all of these sections in the generated linker script at - `arch/x86/kernel/vmlinux.lds` .

That's all. Of course it's not a full description of ELF (Executable and Linkable Format), but if you want to know more, you can find the documentation - here

Inline assembly

Introduction

While reading source code in the Linux kernel, I often see statements like this:

```
__asm__("andq %%rsp,%0; ":"=r" (ti) : "0" (CURRENT_MASK));
```

Yes, this is inline assembly or in other words assembler code which is integrated in a high level programming language. In this case the high level programming language is C. Yes, the `C` programming language is not very high-level, but still.

If you are familiar with the assembly programming language, you may notice that `inline assembly` is not very different from normal assembler. Moreover, the special form of inline assembly which is called `basic form` is exactly the same. For example:

```
__asm__("movq %rax, %rsp");
```

or:

```
__asm__("hlt");
```

The same code (of course without `__asm__` prefix) you might see in plain assembly code. Yes, this is very similar, but not so simple as it might seem at first glance. Actually, the GCC supports two forms of inline assembly statements:

- `basic` ;
- `extended` .

The basic form consists of only two things: the `__asm__` keyword and the string with valid assembler instructions. For example it may look something like this:

```
__asm__("movq    $3, %rax\t\n"
        "movq    %rsi, %rdi");
```

The `asm` keyword may be used in place of `__asm__` , however `__asm__` is portable whereas the `asm` keyword is a `GNU` extension. In further examples I will only use the `__asm__` variant.

If you know assembly programming language this looks pretty familiar. The main problem is in the second form of inline assembly statements - `extended` . This form allows us to pass parameters to an assembly statement, perform jumps etc. Does not sound difficult, but requires knowledge of special rules in addition to knowledge of the assembly language. Every time I see yet another piece of inline assembly code in the Linux kernel, I need to refer to the official documentation of `GCC` to remember how a particular `qualifier` behaves or what the meaning of `=&r` is for example.

I've decided to write this part to consolidate my knowledge related to the inline assembly, as inline assembly statements are quite common in the Linux kernel and we may see them in linux-insides parts sometimes. I thought that it would be useful if we have a special part which contains information on more important aspects of the inline assembly. Of course you may find comprehensive information about inline assembly in the official documentation, but I like to put everything in one place.

Note: This part will not provide guide for assembly programming. It is not intended to teach you to write programs with assembler or to know what one or another assembler instruction means. Just a little memo for extended asm.

Introduction to extended inline assembly

So, let's start. As I already mentioned above, the `basic` assembly statement consists of the `asm` or `__asm__` keyword and set of assembly instructions. This form is in no way different from "normal" assembly. The most interesting part is inline assembler with operands, or `extended` assembler. An extended assembly statement looks more complicated and consists of more than two parts:

```
__asm__ [volatile] [goto] (AssemblerTemplate
                           [ : OutputOperands ]
                           [ : InputOperands  ]
                           [ : Clobbers       ]
                           [ : GotoLabels     ]);
```

All parameters which are marked with squared brackets are optional. You may notice that if we skip the optional parameters and the modifiers `volatile` and `goto` we obtain the `basic` form.

Let's start to consider this in order. The first optional `qualifier` is `volatile`. This specifier tells the compiler that an assembly statement may produce `side effects`. In this case we need to prevent compiler optimizations related to the given assembly statement. In simple terms the `volatile` specifier instructs the compiler not to modify the statement and place it exactly where it was in the original code. As an example let's look at the following function from the Linux kernel:

```
static inline void native_load_gdt(const struct desc_ptr *dtr)
{
    asm volatile("lgdt %0"::"m" (*dtr));
}
```

Here we see the `native_load_gdt` function which loads a base address from the Global Descriptor Table to the `GDTR` register with the `lgdt` instruction. This assembly statement is marked with `volatile` qualifier. It is very important that the compiler does not change the original place of this assembly statement in the resulting code. Otherwise the `GDTR` register may contain wrong address for the `Global Descriptor Table` or the address may be correct, but the structure has not been filled yet. This can lead to an exception being generated, preventing the kernel from booting correctly.

The second optional `qualifier` is the `goto`. This qualifier tells the compiler that the given assembly statement may perform a jump to one of the labels which are listed in the `GotoLabels`. For example:

```
__asm__ goto("jmp %l[label]" : : : : label);
```

Since we finished with these two qualifiers, let's look at the main part of an assembly statement body. As we have seen above, the main part of an assembly statement consists of the following four parts:

- set of assembly instructions;
- output parameters;
- input parameters;
- clobbers.

The first represents a string which contains a set of valid assembly instructions which may be separated by the `\t\n` sequence. Names of processor registers must be prefixed with the `%%` sequence in `extended` form and other symbols like immediates must start with the `$` symbol. The `OutputOperands` and `InputOperands` are comma-separated lists of C variables which may be provided with "constraints" and the `Clobbers` is a list of registers or other values which are modified by the assembler instructions from the `AssemblerTemplate` beyond those listed in the `OutputOperands`. Before we dive into the examples we have to know a little bit about `constraints`. A constraint is a string which specifies placement of an operand. For example the value of an operand may be written to a processor register or read from memory etc.

Consider the following simple example:

```
#include <stdio.h>
```

```
int main(void)
{
        unsigned long a = 5;
        unsigned long b = 10;
        unsigned long sum = 0;

        __asm__("addq %1,%2" : "=r" (sum) : "r" (a), "0" (b));
        printf("a + b = %lu\n", sum);
        return 0;
}
```

Let's compile and run it to be sure that it works as expected:

```
$ gcc test.c -o test
./test
a + b = 15
```

Ok, great. It works. Now let's look at this example in detail. Here we see a simple `C` program which calculates the sum of two variables placing the result into the `sum` variable and in the end we print the result. This example consists of three parts. The first is the assembly statement with the add instruction. It adds the value of the source operand together with the value of the destination operand and stores the result in the destination operand. In our case:

```
addq %1, %2
```

will be expanded to the:

```
addq a, b
```

Variables and expressions which are listed in the `OutputOperands` and `InputOperands` may be matched in the `AssemblerTemplate`. An input/output operand is designated as `%N` where the `N` is the number of operand from left to right beginning from `zero`. The second part of the our assembly statement is located after the first `:` symbol and contains the definition of the output value:

```
"=r" (sum)
```

Notice that the `sum` is marked with two special symbols: `=r`. This is the first constraint that we have encountered. The actual constraint here is only `r` itself. The `=` symbol is `modifier` which denotes output value. This tells to compiler that the previous value will be discarded and replaced by the new data. Besides the `=` modifier, `GCC` provides support for following three modifiers:

- `+` - an operand is read and written by an instruction;
- `&` - output register shouldn't overlap an input register and should be used only for output;
- `%` - tells the compiler that operands may be commutative.

Now let's go back to the `r` qualifier. As I mentioned above, a qualifier denotes the placement of an operand. The `r` symbol means a value will be stored in one of the general purpose register. The last part of our assembly statement:

```
"r" (a), "0" (b)
```

These are input operands - variables `a` and `b`. We already know what the `r` qualifier does. Now we can have a look at the constraint for the variable `b`. The `0` or any other digit from `1` to `9` is called "matching constraint". With this a single operand can be used for multiple roles. The value of the constraint is the source operand index. In our case `0` will match `sum`. If we look at assembly output of our program:

```
0000000000400400 <main>:
  ...
  ...
  ...
  4004fe:       48 c7 45 f8 05 00 00    movq   $0x5,-0x8(%rbp)
  400506:       48 c7 45 f0 0a 00 00    movq   $0xa,-0x10(%rbp)

  400516:       48 8b 55 f8             mov    -0x8(%rbp),%rdx
  40051a:       48 8b 45 f0             mov    -0x10(%rbp),%rax
  40051e:       48 01 d0                add    %rdx,%rax
```

First of all our values `5` and `10` will be put at the stack and then these values will be moved to the two general purpose registers: `%rdx` and `%rax`.

This way the `%rax` register is used for storing the value of the `b` as well as storing the result of the calculation. **NOTE** that I've used `gcc 6.3.1` version, so the resulted code of your compiler may differ.

We have looked at input and output parameters of an inline assembly statement. Before we move on to other constraints supported by `gcc`, there is one remaining part of the inline assembly statement we have not discussed yet - `clobbers`.

Clobbers

As mentioned above, the "clobbered" part should contain a comma-separated list of registers whose content will be modified by the assembler code. This is useful if our assembly expression needs additional registers for calculation. If we add clobbered registers to the inline assembly statement, the compiler take this into account and the register in question will not simultaneously be used by the compiler.

Consider the example from before, but we will add an additional, simple assembler instruction:

```
__asm__("movq $100, %%rdx\t\n"
        "addq %1,%2" : "=r" (sum) : "r" (a), "0" (b));
```

If we look at the assembly output:

```
0000000000400400 <main>:
  ...
  ...
  ...
  4004fe:       48 c7 45 f8 05 00 00    movq   $0x5,-0x8(%rbp)
  400506:       48 c7 45 f0 0a 00 00    movq   $0xa,-0x10(%rbp)

  400516:       48 8b 55 f8             mov    -0x8(%rbp),%rdx
  40051a:       48 8b 45 f0             mov    -0x10(%rbp),%rax

  40051e:       48 c7 c2 64 00 00 00    mov    $0x64,%rdx
  400525:       48 01 d0                add    %rdx,%rax
```

we will see that the `%rdx` register is overwritten with `0x64` or `100` and the result will be `110` instead of `10`. Now if we add the `%rdx` register to the list of `clobbered` registers:

```
__asm__("movq $100, %%rdx\t\n"
        "addq %1,%2" : "=r" (sum) : "r" (a), "0" (b) : "%rdx");
```

and look at the assembler output again:

```
0000000000400400 <main>:
  4004fe:       48 c7 45 f8 05 00 00    movq   $0x5,-0x8(%rbp)
```

```
  400506:       48 c7 45 f0 0a 00 00    movq   $0xa,-0x10(%rbp)

  400516:       48 8b 4d f8             mov    -0x8(%rbp),%rcx
  40051a:       48 8b 45 f0             mov    -0x10(%rbp),%rax

  40051e:       48 c7 c2 64 00 00 00    mov    $0x64,%rdx
  400525:       48 01 c8                add    %rcx,%rax
```

the `%rcx` register will be used for `sum` calculation, preserving the intended semantics of the program. Besides general purpose registers, we may pass two special specifiers. They are:

- `cc` ;
- `memory` .

The first - `cc` indicates that an assembler code modifies flags register. This is typically used if the assembly within contains arithmetic or logic instructions:

```
__asm__("incq %0" ::""(variable): "cc");
```

The second `memory` specifier tells the compiler that the given inline assembly statement executes read/write operations on memory not specified by operands in the output list. This prevents the compiler from keeping memory values loaded and cached in registers. Let's take a look at the following example:

```
#include <stdio.h>

int main(void)
{
        unsigned long a[3] = {10000000000, 0, 1};
        unsigned long b = 5;

        __asm__ volatile("incq %0" :: "m" (a[0]));

        printf("a[0] - b = %lu\n", a[0] - b);
        return 0;
}
```

This example may be artificial, but it illustrates the main idea. Here we have an array of integers and one integer variable. The example is pretty simple, we take the first element of `a` and increment its value. After this we subtract the value of `b` from the first element of `a` . In the end we print the result. If we compile and run this simple example the result may surprise you:

```
~$ gcc -O3  test.c -o test
~$ ./test
a[0] - b = 9999999995
```

The result is `a[0] - b = 9999999995` here, but why? We incremented `a[0]` and subtracted `b` , so the result should be `a[0] - b = 9999999996` here.

If we have a look at the assembler output for this example:

```
00000000004004f6 <main>:
  4004b4:       48 b8 00 e4 0b 54 02    movabs $0x2540be400,%rax
  4004be:       48 89 04 24             mov    %rax,(%rsp)
  ...
  ...
  ...
  40050e:       ff 44 24 f0             incq   (%rsp)

  4004d8:       48 be fb e3 0b 54 02    movabs $0x2540be3fb,%rsi
```

we will see that the first element of the `a` contains the value `0x2540be400` (`10000000000`). The last two lines of code are the actual calculations.

We see our increment instruction with `incq` but then just a move of `0x2540be3fb` (`9999999995`) to the `%rsi` register. This looks strange.

The problem is we have passed the `-O3` flag to `gcc`, so the compiler did some constant folding and propagation to determine the result of `a[0] - 5` at compile time and reduced it to a `movabs` with a constant `0x2540be3fb` or `9999999995` in runtime.

Let's now add `memory` to the clobbers list:

```
__asm__ volatile("incq %0" :: "m" (a[0]) : "memory");
```

and the new result of running this is:

```
~$ gcc -O3  test.c -o test
~$ ./test
a[0] - b = 9999999996
```

Now the result is correct. If we look at the assembly output again:

```
00000000004004f6 <main>:
  400404:       48 b8 00 e4 0b 54 02    movabs $0x2540be400,%rax
  40040b:       00 00 00
  40040e:       48 89 04 24             mov    %rax,(%rsp)
  400412:       48 c7 44 24 08 00 00    movq   $0x0,0x8(%rsp)
  400419:       00 00
  40041b:       48 c7 44 24 10 01 00    movq   $0x1,0x10(%rsp)
  400422:       00 00
  400424:       48 ff 04 24             incq   (%rsp)
  400428:       48 8b 04 24             mov    (%rsp),%rax
  400431:       48 8d 70 fb             lea    -0x5(%rax),%rsi
```

we will see one difference here which is in the last two lines:

```
  400428:       48 8b 04 24             mov    (%rsp),%rax
  400431:       48 8d 70 fb             lea    -0x5(%rax),%rsi
```

Instead of constant folding, `GCC` now preserves calculations in the assembly and places the value of `a[0]` in the `%rax` register afterwards. In the end it just subtracts the constant value of `b` from the `%rax` register and puts result to the `%rsi` .

Besides the `memory` specifier, we also see a new constraint here - `m` . This constraint tells the compiler to use the address of `a[0]` , instead of its value. So, now we are finished with `clobbers` and we may continue by looking at other constraints supported by `GCC` besides `r` and `m` which we have already seen.

Constraints

Now that we are finished with all three parts of an inline assembly statement, let's return to constraints. We already saw some constraints in the previous parts, like `r` which represents a `register` operand, `m` which represents a memory operand and `0-9` which represent an reused, indexed operand. Besides these `GCC` provides support for other constraints. For example the `i` constraint represents an `immediate` integer operand with know value:

```
#include <stdio.h>

int main(void)
{
```

```
        int a = 0;

        __asm__("movl %1, %0" : "=r"(a) : "i"(100));
        printf("a = %d\n", a);
        return 0;
}
```

The result is:

```
~$ gcc test.c -o test
~$ ./test
a = 100
```

Or for example `I` which represents an immediate 32-bit integer. The difference between `i` and `I` is that `i` is general, whereas `I` is strictly specified to 32-bit integer data. For example if you try to compile the following code:

```
unsigned long test_asm(int nr)
{
        unsigned long a = 0;

        __asm__("movq %1, %0" : "=r"(a) : "I"(0xffffffffffff));
        return a;
}
```

you will get an error:

```
$ gcc -O3 test.c -o test
test.c: In function ‘test_asm’:
test.c:7:9: warning: asm operand 1 probably doesn’t match constraints
         __asm__("movq %1, %0" : "=r"(a) : "I"(0xffffffffffff));
         ^
test.c:7:9: error: impossible constraint in ‘asm’
```

when at the same time:

```
unsigned long test_asm(int nr)
{
        unsigned long a = 0;

        __asm__("movq %1, %0" : "=r"(a) : "i"(0xffffffffffff));
        return a;
}
```

works perfectly:

```
~$ gcc -O3 test.c -o test
~$ echo $?
0
```

`GCC` also supports `J`, `K`, `N` constraints for integer constants in the range of 0-63 bits, signed 8-bit integer constants and unsigned 8-bit integer constants respectively. The `o` constraint represents a memory operand with an `offsetable` memory address. For example:

```
#include <stdio.h>

int main(void)
{
        static unsigned long arr[3] = {0, 1, 2};
        static unsigned long element;
```

```
        __asm__ volatile("movq 16+%1, %0" : "=r"(element) : "o"(arr));
        printf("%lu\n", element);
        return 0;
}
```

The result, as expected:

```
~$ gcc -O3 test.c -o test
~$ ./test
2
```

All of these constraints may be combined (so long as they do not conflict). In this case the compiler will choose the best one for a certain situation. For example:

```
unsigned long a = 10;
unsigned long b = 20;

void main(void)
{
    __asm__ ("movq %1,%0" : "=mr"(b) : "rm"(a));
}
```

will use a memory operand:

```
main:
        movq a(%rip),b(%rip)
        ret
b:
        .quad   20
a:
        .quad   10
```

instead of direct usage of general purpose registers.

That's about all of the commonly used constraints in inline assembly statements. You can find more in the official documentation.

Architecture specific constraints

Before we finish, let's look at the set of special constraints. These constrains are architecture specific and as this book is specific to the x86_64 architecture, we will look at constraints related to it. First of all the set of `a` ... `d` and also `S` and `D` constraints represent generic purpose registers. In this case the `a` constraint corresponds to `%al`, `%ax`, `%eax` or `%rax` register depending on instruction size. The `S` and `D` constraints are `%si` and `%di` registers respectively. For example let's take our previous example. We can see in its assembly output that value of the `a` variable is stored in the `%eax` register. Now let's look at the assembly output of the same assembly, but with other constraint:

```
#include <stdio.h>

int a = 1;

int main(void)
{
        int b;
        __asm__ ("movq %1,%0" : "=r"(b) : "d"(a));
        return b;
}
```

Now we see that value of the `a` variable will be stored in the `%rax` register:

```
0000000000400400 <main>:
  4004aa:       48 8b 05 6f 0b 20 00    mov    0x200b6f(%rip),%rax        # 601020 <a>
```

The `f` and `t` constraints represent any floating point stack register - `%st` and the top of the floating point stack respectively. The `u` constraint represents the second value from the top of the floating point stack.

That's all. You may find more details about x86_64 and general constraints in the official documentation.

Links

- Linux kernel source code
- assembly programming language
- GCC
- GNU extension
- Global Descriptor Table
- Processor registers
- add instruction
- flags register
- x86_64
- constraints

Misc

This chapter contains parts which are not directly related to the Linux kernel source code and implementation of different subsystems.

Linux kernel development

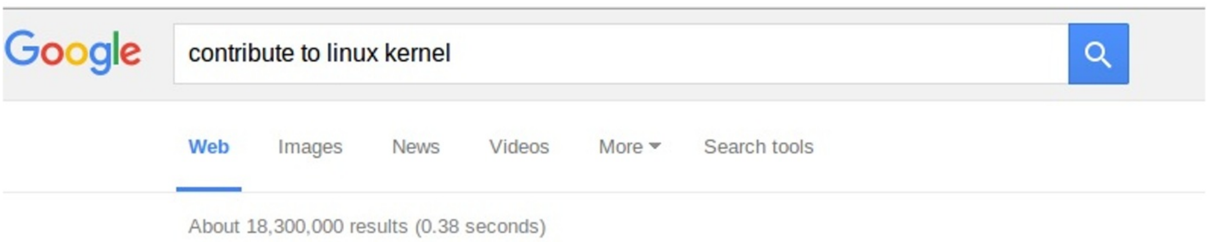
Introduction

As you already may know, I've started a series of blog posts about assembler programming for `x86_64` architecture in the last year. I have never written a line of low-level code before this moment, except for a couple of toy `Hello World` examples in university. It was a long time ago and, as I already said, I didn't write low-level code at all. Some time ago I became interested in such things. I understood that I can write programs, but didn't actually understand how my program is arranged.

After writing some assembler code I began to understand how my program looks after compilation, **approximately**. But anyway, I didn't understand many other things. For example: what occurs when the `syscall` instruction is executed in my assembler, what occurs when the `printf` function starts to work or how can my program talk with other computers via network. Assembler programming language didn't give me answers to my questions and I decided to go deeper in my research. I started to learn from the source code of the Linux kernel and tried to understand the things that I'm interested in. The source code of the Linux kernel didn't give me the answers to **all** of my questions, but now my knowledge about the Linux kernel and the processes around it is much better.

I'm writing this part nine and a half months after I've started to learn from the source code of the Linux kernel and published the first part of this book. Now it contains forty parts and it is not the end. I decided to write this series about the Linux kernel mostly for myself. As you know the Linux kernel is very huge piece of code and it is easy to forget what does this or that part of the Linux kernel mean and how does it implement something. But soon the linux-insides repo became popular and after nine months it has `9096` stars:

It seems that people are interested in the insides of the Linux kernel. Besides this, in all the time that I have been writing `linux-insides`, I have received many questions from different people about how to begin contributing to the Linux kernel. Generally people are interested in contributing to open source projects and the Linux kernel is not an exception:



So, it seems that people are interested in the Linux kernel development process. I thought it would be strange if a book about the Linux kernel would not contain a part describing how to take a part in the Linux kernel development and that's why I decided to write it. You will not find information about why you should be interested in contributing to the Linux kernel in this part. But if you are interested how to start with Linux kernel development, this part is for you.

Let's start.

How to start with Linux kernel

First of all, let's see how to get, build, and run the Linux kernel. You can run your custom build of the Linux kernel in two ways:

- Run the Linux kernel on a virtual machine;
- Run the Linux kernel on real hardware.

I'll provide descriptions for both methods. Before we start doing anything with the Linux kernel, we need to get it. There are a couple of ways to do this depending on your purpose. If you just want to update the current version of the Linux kernel on your computer, you can use the instructions specific to your Linux distro.

In the first case you just need to download new version of the Linux kernel with the package manager. For example, to upgrade the version of the Linux kernel to `4.1` for Ubuntu (Vivid Vervet), you will just need to execute the following commands:

```
$ sudo add-apt-repository ppa:kernel-ppa/ppa
$ sudo apt-get update
```

After this execute this command:

```
$ apt-cache showpkg linux-headers
```

and choose the version of the Linux kernel in which you are interested. In the end execute the next command and replace `${version}` with the version that you chose in the output of the previous command:

```
$ sudo apt-get install linux-headers-${version} linux-headers-${version}-generic linux-image-${version}-generic
 --fix-missing
```

and reboot your system. After the reboot you will see the new kernel in the grub menu.

In the other way if you are interested in the Linux kernel development, you will need to get the source code of the Linux kernel. You can find it on the kernel.org website and download an archive with the Linux kernel source code. Actually the Linux kernel development process is fully built around `git` version control system. So you can get it with `git` from the `kernel.org` :

```
$ git clone git://git.kernel.org/pub/scm/linux/kernel/git/torvalds/linux.git
```

I don't know how about you, but I prefer `github` . There is a mirror of the Linux kernel mainline repository, so you can clone it with:

```
$ git clone git@github.com:torvalds/linux.git
```

I use my own fork for development and when I want to pull updates from the main repository I just execute the following command:

```
$ git checkout master
$ git pull upstream master
```

Note that the remote name of the main repository is `upstream` . To add a new remote with the main Linux repository you can execute:

```
git remote add upstream git@github.com:torvalds/linux.git
```

After this you will have two remotes:

```
~/dev/linux (master) $ git remote -v
origin    git@github.com:0xAX/linux.git (fetch)
origin    git@github.com:0xAX/linux.git (push)
```

```
upstream    https://github.com/torvalds/linux.git (fetch)
upstream    https://github.com/torvalds/linux.git (push)
```

One is of your fork (`origin`) and the second is for the main repository (`upstream`).

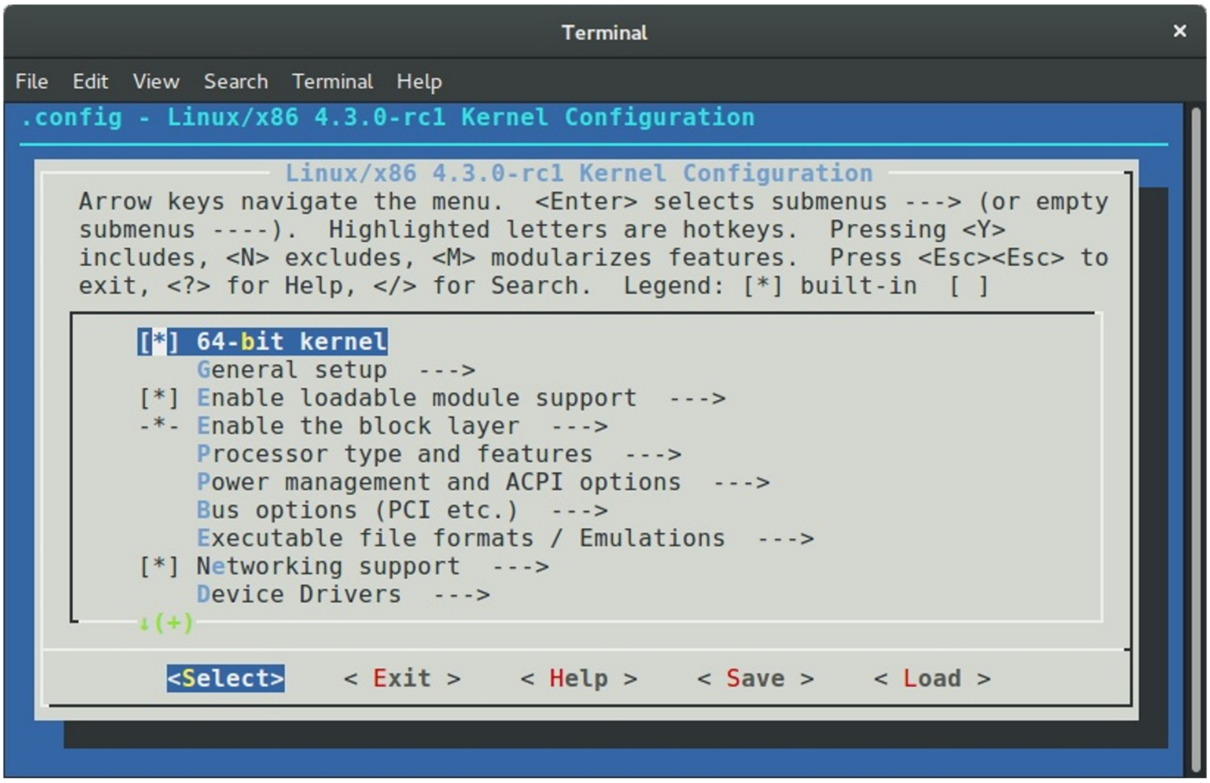
Now that we have a local copy of the Linux kernel source code, we need to configure and build it. The Linux kernel can be configured in different ways. The simplest way is to just copy the configuration file of the already installed kernel that is located in the `/boot` directory:

```
$ sudo cp /boot/config-$(uname -r) ~/dev/linux/.config
```

If your current Linux kernel was built with the support for access to the `/proc/config.gz` file, you can copy your actual kernel configuration file with this command:

```
$ cat /proc/config.gz | gunzip > ~/dev/linux/.config
```

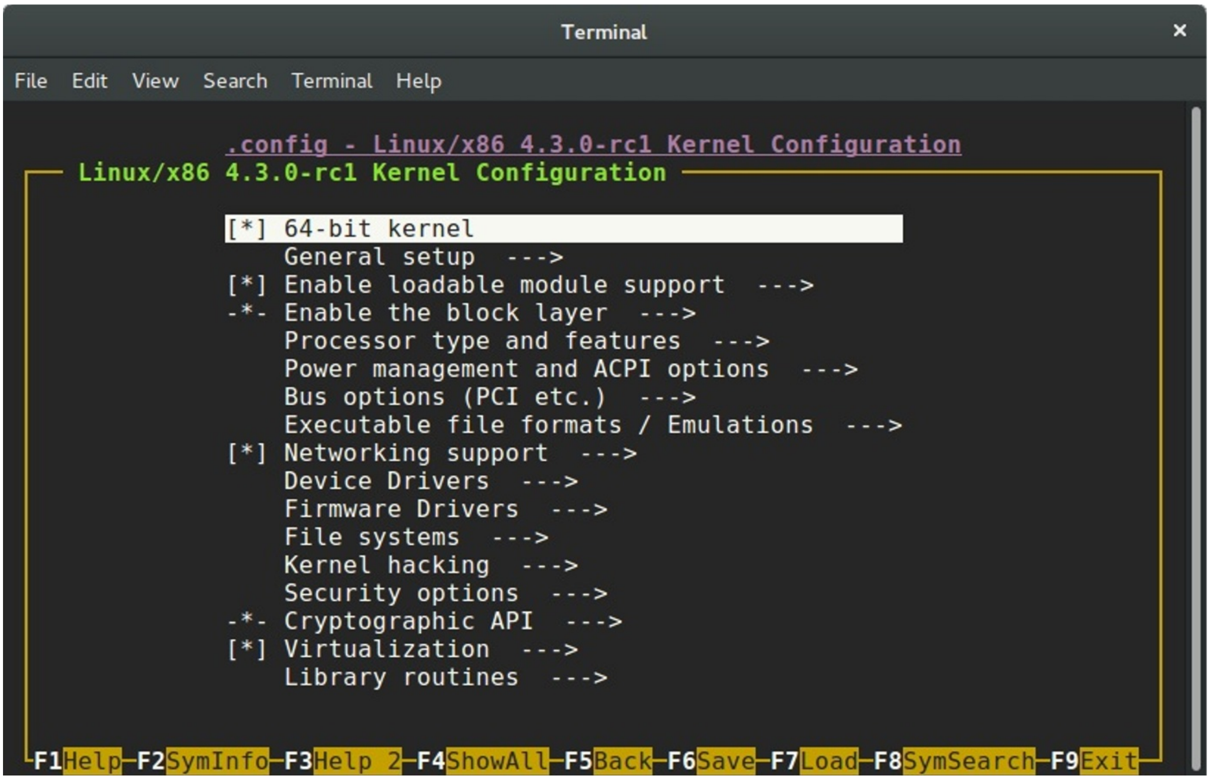
If you are not satisfied with the standard kernel configuration that is provided by the maintainers of your distro, you can configure the Linux kernel manually. There are a couple of ways to do it. The Linux kernel root Makefile provides a set of targets that allows you to configure it. For example `menuconfig` provides a menu-driven interface for the kernel configuration:



The `defconfig` argument generates the default kernel configuration file for the current architecture, for example x86_64 defconfig. You can pass the `ARCH` command line argument to `make` to build `defconfig` for the given architecture:

```
$ make ARCH=arm64 defconfig
```

The `allnoconfig` , `allyesconfig` and `allmodconfig` arguments allow you to generate a new configuration file where all options will be disabled, enabled, and enabled as modules respectively. The `nconfig` command line arguments that provides `ncurses` based program with menu to configure Linux kernel:



And even `randconfig` to generate random Linux kernel configuration file. I will not write about how to configure the Linux kernel or which options to enable because it makes no sense to do so for two reasons: First of all I do not know your hardware and second, if you know your hardware, the only remaining task is to find out how to use programs for kernel configuration, and all of them are pretty simple to use.

OK, we now have the source code of the Linux kernel and configured it. The next step is the compilation of the Linux kernel. The simplest way to compile Linux kernel is to just execute:

```
$ make
scripts/kconfig/conf  --silentoldconfig Kconfig
#
# configuration written to .config
#
  CHK     include/config/kernel.release
  UPD     include/config/kernel.release
  CHK     include/generated/uapi/linux/version.h
  CHK     include/generated/utsrelease.h
  ...
  ...
  ...
  OBJCOPY arch/x86/boot/vmlinux.bin
  AS      arch/x86/boot/header.o
  LD      arch/x86/boot/setup.elf
  OBJCOPY arch/x86/boot/setup.bin
  BUILD   arch/x86/boot/bzImage
  Setup is 15740 bytes (padded to 15872 bytes).
System is 4342 kB
CRC 82703414
Kernel: arch/x86/boot/bzImage is ready  (#73)
```

To increase the speed of kernel compilation you can pass `-jN` command line argument to `make`, where `N` specifies the number of commands to run simultaneously:

```
$ make -j8
```

If you want to build Linux kernel for an architecture that differs from your current, the simplest way to do it pass two arguments:

- `ARCH` command line argument and the name of the target architecture;
- `CROSS_COMPILER` command line argument and the cross-compiler tool prefix;

For example if we want to compile the Linux kernel for the arm64 with default kernel configuration file, we need to execute following command:

```
$ make -j4 ARCH=arm64 CROSS_COMPILER=aarch64-linux-gnu- defconfig
$ make -j4 ARCH=arm64 CROSS_COMPILER=aarch64-linux-gnu-
```

As result of compilation we can see the compressed kernel - `arch/x86/boot/bzImage`. Now that we have compiled the kernel, we can either install it on our computer or just run it in an emulator.

Installing Linux kernel

As I already wrote we will consider two ways how to launch new kernel: In the first case we can install and run the new version of the Linux kernel on the real hardware and the second is launch the Linux kernel on a virtual machine. In the previous paragraph we saw how to build the Linux kernel from source code and as a result we have got compressed image:

```
...
...
...
Kernel: arch/x86/boot/bzImage is ready  (#73)
```

After we have got the bzImage we need to install `headers`, `modules` of the new Linux kernel with the:

```
$ sudo make headers_install
$ sudo make modules_install
```

and directly the kernel itself:

```
$ sudo make install
```

From this moment we have installed new version of the Linux kernel and now we must tell the `bootloader` about it. Of course we can add it manually by the editing of the `/boot/grub2/grub.cfg` configuration file, but I prefer to use a script for this purpose. I'm using two different Linux distros: Fedora and Ubuntu. There are two different ways to update the grub configuration file. I'm using following script for this purpose:

```
#!/bin/bash

source "term-colors"

DISTRIBUTIVE=$(cat /etc/*-release | grep NAME | head -1 | sed -n -e 's/NAME\=//p')
echo -e "Distributive: ${Green}${DISTRIBUTIVE}${Color_Off}"

if [[ "$DISTRIBUTIVE" == "Fedora" ]] ;
then
    su -c 'grub2-mkconfig -o /boot/grub2/grub.cfg'
else
    sudo update-grub
fi

echo "${Green}Done.${Color_Off}"
```

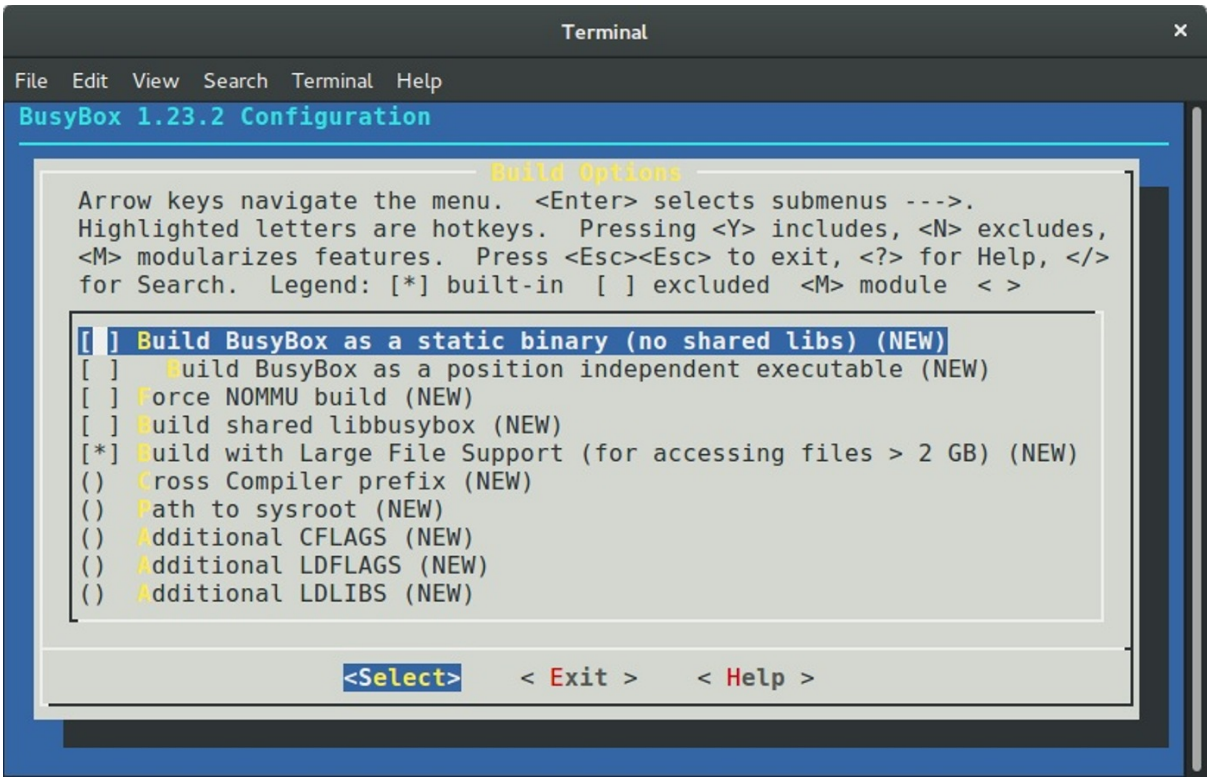
This is the last step of the new Linux kernel installation and after this you can reboot your computer and select new version of the kernel during boot.

The second case is to launch new Linux kernel in the virtual machine. I prefer qemu. First of all we need to build initial ramdisk - initrd for this. The `initrd` is a temporary root file system that is used by the Linux kernel during initialization process while other filesystems are not mounted. We can build `initrd` with the following commands:

First of all we need to download busybox and run `menuconfig` for its configuration:

```
$ mkdir initrd
$ cd initrd
$ curl http://busybox.net/downloads/busybox-1.23.2.tar.bz2 | tar xjf -
$ cd busybox-1.23.2/
$ make menuconfig
$ make -j4
```

`busybox` is an executable file - `/bin/busybox` that contains a set of standard tools like coreutils. In the `busysbox` menu we need to enable: `Build BusyBox as a static binary (no shared libs)` option:



We can find this menu in the:

```
Busybox Settings
--> Build Options
```

After this we exit from the `busysbox` configuration menu and execute following commands for building and installation of it:

```
$ make -j4
$ sudo make install
```

Now that `busybox` is installed, we can begin building our `initrd`. To do this, we go to the previous `initrd` directory and:

```
$ cd ..
$ mkdir -p initramfs
```

```
$ cd initramfs
$ mkdir -pv {bin,sbin,etc,proc,sys,usr/{bin,sbin}}
$ cp -av ../busybox-1.23.2/_install/* .
```

copy `busybox` fields to the `bin` , `sbin` and other directories. Now we need to create executable `init` file that will be executed as a first process in the system. My `init` file just mounts procfs and sysfs filesystems and executed shell:

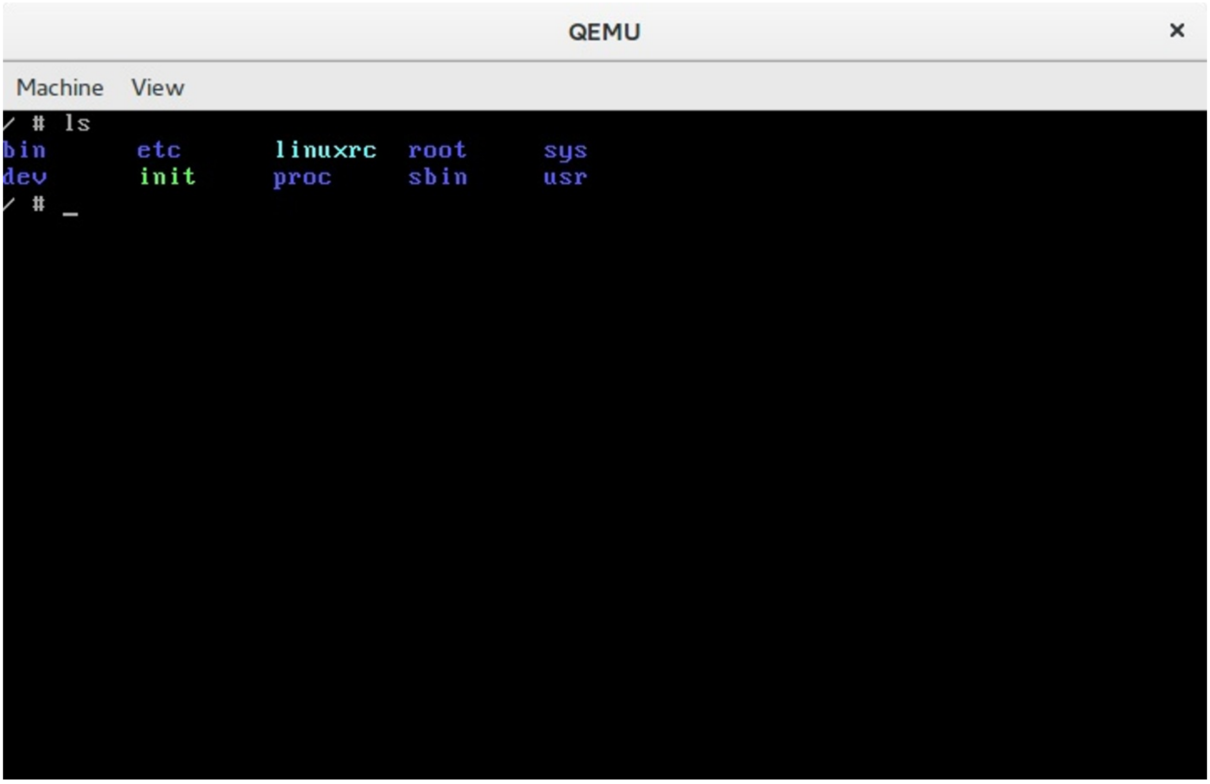
```
#!/bin/sh

mount -t proc none /proc
mount -t sysfs none /sys

exec /bin/sh
```

Now we can create an archive that will be our `initrd` :

```
$ find . -print0 | cpio --null -ov --format=newc | gzip -9 > ~/dev/initrd_x86_64.gz
```

We can now run our kernel in the virtual machine. As I already wrote I prefer qemu for this. We can run our kernel with the following command:

```
$ qemu-system-x86_64 -snapshot -m 8GB -serial stdio -kernel ~/dev/linux/arch/x86_64/boot/bzImage -initrd ~/dev/
initrd_x86_64.gz -append "root=/dev/sda1 ignore_loglevel"
```



From now we can run the Linux kernel in the virtual machine and this means that we can begin to change and test the kernel.

Consider using ivandaviov/minimal or Buildroot to automate the process of generating initrd.

Getting started with the Linux Kernel Development

The main point of this paragraph is to answer two questions: What to do and what not to do before sending your first patch to the Linux kernel. Please, do not confuse this `to do` with `todo`. I have no answer what you can fix in the Linux kernel. I just want to tell you my workflow during experimenting with the Linux kernel source code.

First of all I pull the latest updates from Linus's repo with the following commands:

```
$ git checkout master
$ git pull upstream master
```

After this my local repository with the Linux kernel source code is synced with the mainline repository. Now we can make some changes in the source code. As I already wrote, I have no advice for you where you can start and what `TODO` in the Linux kernel. But the best place for newbies is `staging` tree. In other words the set of drivers from the drivers/staging. The maintainer of the `staging` tree is Greg Kroah-Hartman and the `staging` tree is that place where your trivial patch can be accepted. Let's look on a simple example that describes how to generate patch, check it and send to the Linux kernel mail listing.

If we look in the driver for the Digi International EPCA PCI based devices, we will see the `dgap_sindex` function on line 295:

```
static char *dgap_sindex(char *string, char *group)
{
    char *ptr;

    if (!string || !group)
        return NULL;

    for (; *string; string++) {
        for (ptr = group; *ptr; ptr++) {
            if (*ptr == *string)
                return string;
        }
    }

    return NULL;
}
```

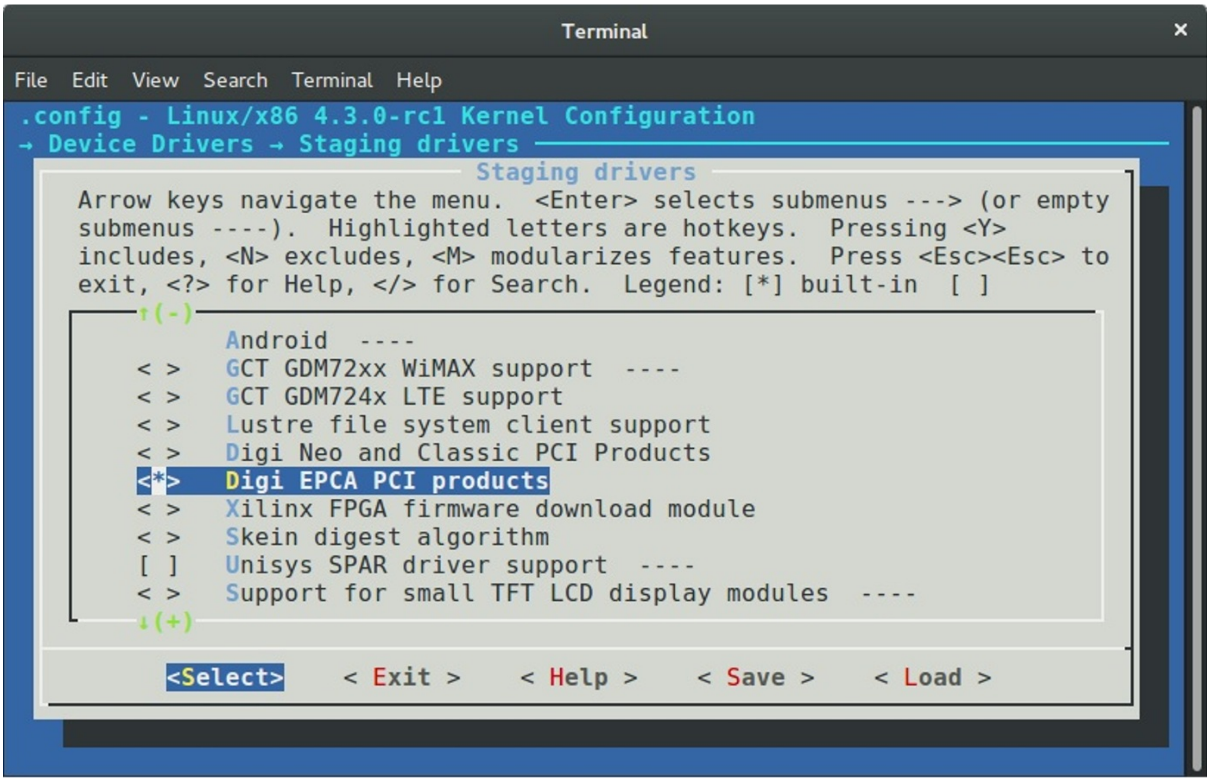
This function looks for a match of any character in the group and returns that position. During research of source code of the Linux kernel, I have noted that the lib/string.c source code file contains the implementation of the `strpbrk` function that does the same thing as `dgap_sinidex`. It is not a good idea to use a custom implementation of a function that already exists, so we can remove the `dgap_sindex` function from the drivers/staging/dgap/dgap.c source code file and use the `strpbrk` instead.

First of all let's create new `git` branch based on the current master that synced with the Linux kernel mainline repo:

```
$ git checkout -b "dgap-remove-dgap_sindex"
```

And now we can replace the `dgap_sindex` with the `strpbrk`. After we did all changes we need to recompile the Linux kernel or just dgap directory. Do not forget to enable this driver in the kernel configuration. You can find it in the:

```
Device Drivers
--> Staging drivers
----> Digi EPCA PCI products
```



Now is time to make commit. I'm using following combination for this:

```
$ git add .
$ git commit -s -v
```

After the last command an editor will be opened that will be chosen from `$GIT_EDITOR` or `$EDITOR` environment variable. The `-s` command line argument will add `Signed-off-by` line by the committer at the end of the commit log message. You can find this line in the end of each commit message, for example - 00cc1633. The main point of this line is the tracking of who did a change. The `-v` option show unified diff between the HEAD commit and what would be committed at the bottom of the commit message. It is not necessary, but very useful sometimes. A couple of words about commit message. Actually a commit message consists from two parts:

The first part is on the first line and contains short description of changes. It starts from the `[PATCH]` prefix followed by a subsystem, driver or architecture name and after `:` symbol short description. In our case it will be something like this:

```
[PATCH] staging/dgap: Use strpbrk() instead of dgap_sindex()
```

After short description usually we have an empty line and full description of the commit. In our case it will be:

```
The <linux/string.h> provides strpbrk() function that does the same that the
dgap_sindex(). Let's use already defined function instead of writing custom.
```

And the `Sign-off-by` line in the end of the commit message. Note that each line of a commit message must no be longer than `80` symbols and commit message must describe your changes in details. Do not just write a commit message like: `Custom function removed`, you need to describe what you did and why. The patch reviewers must know what they review. Besides this commit messages in this view are very helpful. Each time when we can't understand something, we can use git blame to read description of changes.

After we have committed changes time to generate patch. We can do it with the `format-patch` command:

```
$ git format-patch master
0001-staging-dgap-Use-strpbrk-instead-of-dgap_sindex.patch
```

We've passed name of the branch (`master` in this case) to the `format-patch` command that will generate a patch with the last changes that are in the `dgap-remove-dgap_sindex` branch and not are in the `master` branch. As you can note, the `format-patch` command generates file that contains last changes and has name that is based on the commit short description. If you want to generate a patch with the custom name, you can use `--stdout` option:

```
$ git format-patch master --stdout > dgap-patch-1.patch
```

The last step after we have generated our patch is to send it to the Linux kernel mailing list. Of course, you can use any email client, `git` provides a special command for this: `git send-email` . Before you send your patch, you need to know where to send it. Yes, you can just send it to the Linux kernel mailing list address which is `linux-kernel@vger.kernel.org` , but it is very likely that the patch will be ignored, because of the large flow of messages. The better choice would be to send the patch to the maintainers of the subsystem where you have made changes. To find the names of these maintainers use the `get_maintainer.pl` script. All you need to do is pass the file or directory where you wrote code.

```
$ ./scripts/get_maintainer.pl -f drivers/staging/dgap/dgap.c
Lidza Louina <lidza.louina@gmail.com> (maintainer:DIGI EPCA PCI PRODUCTS)
Mark Hounschell <markh@compro.net> (maintainer:DIGI EPCA PCI PRODUCTS)
Daeseok Youn <daeseok.youn@gmail.com> (maintainer:DIGI EPCA PCI PRODUCTS)
Greg Kroah-Hartman <gregkh@linuxfoundation.org> (supporter:STAGING SUBSYSTEM)
driverdev-devel@linuxdriverproject.org (open list:DIGI EPCA PCI PRODUCTS)
devel@driverdev.osuosl.org (open list:STAGING SUBSYSTEM)
linux-kernel@vger.kernel.org (open list)
```

You will see the set of the names and related emails. Now we can send our patch with:

```
$ git send-email --to "Lidza Louina <lidza.louina@gmail.com>" \
  --cc "Mark Hounschell <markh@compro.net>"                   \
  --cc "Daeseok Youn <daeseok.youn@gmail.com>"                \
  --cc "Greg Kroah-Hartman <gregkh@linuxfoundation.org>"      \
  --cc "driverdev-devel@linuxdriverproject.org"               \
  --cc "devel@driverdev.osuosl.org"                           \
  --cc "linux-kernel@vger.kernel.org"
```

That's all. The patch is sent and now you only have to wait for feedback from the Linux kernel developers. After you send a patch and a maintainer accepts it, you will find it in the maintainer's repository (for example patch that you saw in this part) and after some time the maintainer will send a pull request to Linus and you will see your patch in the mainline repository.

That's all.

Some advice

In the end of this part I want to give you some advice that will describe what to do and what not to do during development of the Linux kernel:

- Think, Think, Think. And think again before you decide to send a patch.
- Each time when you have changed something in the Linux kernel source code - compile it. After any changes. Again and again. Nobody likes changes that don't even compile.
- The Linux kernel has a coding style guide and you need to comply with it. There is great script which can help to check your changes. This script is - scripts/checkpatch.pl. Just pass source code file with changes to it and you will see:

```
$ ./scripts/checkpatch.pl -f drivers/staging/dgap/dgap.c
WARNING: Block comments use * on subsequent lines
#94: FILE: drivers/staging/dgap/dgap.c:94:
+/*
+     SUPPORTED PRODUCTS

CHECK: spaces preferred around that '|' (ctx:VxV)
#143: FILE: drivers/staging/dgap/dgap.c:143:
+    { PPCM,        PCI_DEV_XEM_NAME,     64, (T_PCXM|T_PCLITE|T_PCIBUS) },
```

Also you can see problematic places with the help of the `git diff` :

```
~/dev/linux (dgap-remove-dgap_sindex) $ git diff
diff --git a/init/main.c b/init/main.c
index 9e64d70..af379a5 100644
--- a/init/main.c
+++ b/init/main.c
@@ -153,6 +153,8 @@ EXPORT_SYMBOL(reset_devices);
 static int __init set_reset_devices(char *str)
 {
        reset_devices = 1;
+
+
        return 1;
 }
```

- Linus doesn't accept github pull requests

- If your change consists from some different and unrelated changes, you need to split the changes via separate commits. The `git format-patch` command will generate patches for each commit and the subject of each patch will contain a `vN` prefix where the `N` is the number of the patch. If you are planning to send a series of patches it will be helpful to pass the `--cover-letter` option to the `git format-patch` command. This will generate an additional file that will contain the cover letter that you can use to describe what your patchset changes. It is also a good idea to use the `--in-reply-to` option in the `git send-email` command. This option allows you to send your patch series in reply to your cover message. The structure of the your patch will look like this for a maintainer:

```
|--> cover letter
  |----> patch_1
  |----> patch_2
```

You need to pass `message-id` as an argument of the `--in-reply-to` option that you can find in the output of the `git send-email` :

It's important that your email be in the plain text format. Generally, `send-email` and `format-patch` are very useful during development, so look at the documentation for the commands and you'll find some useful options such as: git send-email and git format-patch.

- Do not be surprised if you do not get an immediate answer after you send your patch. Maintainers can be very busy.

- The scripts directory contains many different useful scripts that are related to Linux kernel development. We already saw two scripts from this directory: the `checkpatch.pl` and the `get_maintainer.pl` scripts. Outside of those scripts, you can find the stackusage script that will print usage of the stack, extract-vmlinux for extracting an uncompressed kernel image, and many others. Outside of the `scripts` directory you can find some very useful scripts by Lorenzo Stoakes for kernel development.

* Subscribe to the Linux kernel mailing list. There are a large number of letters every day on `lkml`, but it is very useful to read them and understand things such as the current state of the Linux kernel. Other than `lkml` there are set mailing listings which are related to the different Linux kernel subsystems.
* If your patch is not accepted the first time and you receive feedback from Linux kernel developers, make your changes and resend the patch with the `[PATCH vN]` prefix (where `N` is the number of patch version). For example:

```
[PATCH v2] staging/dgap: Use strpbrk() instead of dgap_sindex()
```

Also it must contain a changelog that describes all changes from previous patch versions. Of course, this is not an exhaustive list of requirements for Linux kernel development, but some of the most important items were addressed.

Happy Hacking!

Conclusion

I hope this will help others join the Linux kernel community! If you have any questions or suggestions, write me at email or ping me on twitter.

Please note that English is not my first language, and I am really sorry for any inconvenience. If you find any mistakes please let me know via email or send a PR.

Links

* blog posts about assembly programming for x86_64
* Assembler
* distro
* package manager
* grub
* kernel.org
* version control system
* arm64
* bzImage
* qemu
* initrd
* busybox
* coreutils
* procfs
* sysfs
* Linux kernel mail listing archive
* Linux kernel coding style guide
* How to Get Your Change Into the Linux Kernel
* Linux Kernel Newbies
* plain text

Process of the Linux kernel building

Introduction

I won't tell you how to build and install a custom Linux kernel on your machine. If you need help with this, you can find many resources that will help you do it. Instead, we will learn what occurs when you execute `make` in the root directory of the Linux kernel source code.

When I started to study the source code of the Linux kernel, the makefile was the first file that I opened. And it was scary :). The makefile contained `1591` lines of code when I wrote this part and the kernel was the 4.2.0-rc3 release.

This makefile is the top makefile in the Linux kernel source code and the kernel building starts here. Yes, it is big, but moreover, if you've read the source code of the Linux kernel you may have noted that all directories containing source code has its own makefile. Of course it is not possible to describe how each source file is compiled and linked, so we will only study the standard compilation case. You will not find here building of the kernel's documentation, cleaning of the kernel source code, tags generation, cross-compilation related stuff, etc... We will start from the `make` execution with the standard kernel configuration file and will finish with the building of the bzImage.

It would be better if you're already familiar with the make util, but I will try to describe every piece of code in this part anyway.

So let's start.

Preparation before the kernel compilation

There are many things to prepare before the kernel compilation can be started. The main point here is to find and configure the type of compilation, to parse command line arguments that are passed to `make`, etc... So let's dive into the top `Makefile` of Linux kernel.

The top `Makefile` of Linux kernel is responsible for building two major products: vmlinux (the resident kernel image) and the modules (any module files). The Makefile of the Linux kernel starts with the definition of following variables:

```
VERSION = 4
PATCHLEVEL = 2
SUBLEVEL = 0
EXTRAVERSION = -rc3
NAME = Hurr durr I'ma sheep
```

These variables determine the current version of Linux kernel and are used in different places, for example in the forming of the `KERNELVERSION` variable in the same `Makefile` :

```
KERNELVERSION = $(VERSION)$(if $(PATCHLEVEL),.$(PATCHLEVEL)$(if $(SUBLEVEL),.$(SUBLEVEL)))$(EXTRAVERSION)
```

After this we can see a couple of `ifeq` conditions that check some of the parameters passed to `make` . The Linux kernel `makefiles` provides a special `make help` target that prints all available targets and some of the command line arguments that can be passed to `make` . For example : `make V=1` => verbose build. The first `ifeq` checks whether the `V=n` option is passed to `make` :

```
ifeq ("$(origin V)", "command line")
  KBUILD_VERBOSE = $(V)
endif
ifndef KBUILD_VERBOSE
  KBUILD_VERBOSE = 0
```

```
endif

ifeq ($(KBUILD_VERBOSE),1)
  quiet =
  Q =
else
  quiet=quiet_
  Q = @
endif

export quiet Q KBUILD_VERBOSE
```

If this option is passed to `make`, we set the `KBUILD_VERBOSE` variable to the value of `V` option. Otherwise we set the `KBUILD_VERBOSE` variable to zero. After this we check the value of `KBUILD_VERBOSE` variable and set values of the `quiet` and `Q` variables depending on the value of `KBUILD_VERBOSE` variable. The `@` symbols suppress the output of command. And if it is present before a command the output will be something like this: `CC scripts/mod/empty.o` instead of `Compiling .... scripts/mod/empty.o`. In the end we just export all of these variables. The next `ifeq` statement checks that `O=/dir` option was passed to the `make`. This option allows to locate all output files in the given `dir`:

```
ifeq ($(KBUILD_SRC),)

ifeq ("$(origin O)", "command line")
  KBUILD_OUTPUT := $(O)
endif

ifneq ($(KBUILD_OUTPUT),)
saved-output := $(KBUILD_OUTPUT)
KBUILD_OUTPUT := $(shell mkdir -p $(KBUILD_OUTPUT) && cd $(KBUILD_OUTPUT) \
                                && /bin/pwd)
$(if $(KBUILD_OUTPUT),, \
     $(error failed to create output directory "$(saved-output)"))

sub-make: FORCE
    $(Q)$(MAKE) -C $(KBUILD_OUTPUT) KBUILD_SRC=$(CURDIR) \
    -f $(CURDIR)/Makefile $(filter-out _all sub-make,$(MAKECMDGOALS))

skip-makefile := 1
endif # ifneq ($(KBUILD_OUTPUT),)
endif # ifeq ($(KBUILD_SRC),)
```

We check the `KBUILD_SRC` that represents the top directory of the kernel source code and whether it is empty (it is empty when the makefile is executed for the first time). We then set the `KBUILD_OUTPUT` variable to the value passed with the `O` option (if this option was passed). In the next step we check this `KBUILD_OUTPUT` variable and if it is set, we do following things:

- Store the value of `KBUILD_OUTPUT` in the temporary `saved-output` variable;
- Try to create the given output directory;
- Check that directory created, in other way print error message;
- If the custom output directory was created successfully, execute `make` again with the new directory (see the `-C` option).

The next `ifeq` statements check that the `C` or `M` options passed to `make`:

```
ifeq ("$(origin C)", "command line")
  KBUILD_CHECKSRC = $(C)
endif
ifndef KBUILD_CHECKSRC
  KBUILD_CHECKSRC = 0
endif

ifeq ("$(origin M)", "command line")
  KBUILD_EXTMOD := $(M)
endif
```

The `C` option tells the `makefile` that we need to check all `c` source code with a tool provided by the `$CHECK` environment variable, by default it is sparse. The second `M` option provides build for the external modules (will not see this case in this part). We also check whether the `KBUILD_SRC` variable is set, and if it isn't, we set the `srctree` variable to `.`:

```
ifeq ($(KBUILD_SRC),)
        srctree := .
endif

objtree    := .
src        := $(srctree)
obj        := $(objtree)

export srctree objtree VPATH
```

That tells `Makefile` that the kernel source tree will be in the current directory where `make` was executed. We then set `objtree` and other variables to this directory and export them. The next step is to get value for the `SUBARCH` variable that represents what the underlying architecture is:

```
SUBARCH := $(shell uname -m | sed -e s/i.86/x86/ -e s/x86_64/x86/ \
                  -e s/sun4u/sparc64/ \
                  -e s/arm.*/arm/ -e s/sa110/arm/ \
                  -e s/s390x/s390/ -e s/parisc64/parisc/ \
                  -e s/ppc.*/powerpc/ -e s/mips.*/mips/ \
                  -e s/sh[234].*/sh/ -e s/aarch64.*/arm64/ )
```

As you can see, it executes the uname util that prints information about machine, operating system and architecture. As it gets the output of `uname`, it parses the output and assigns the result to the `SUBARCH` variable. Now that we have `SUBARCH`, we set the `SRCARCH` variable that provides the directory of the certain architecture and `hfr-arch` that provides the directory for the header files:

```
ifeq ($(ARCH),i386)
        SRCARCH := x86
endif
ifeq ($(ARCH),x86_64)
        SRCARCH := x86
endif

hdr-arch  := $(SRCARCH)
```

Note `ARCH` is an alias for `SUBARCH`. In the next step we set the `KCONFIG_CONFIG` variable that represents path to the kernel configuration file and if it was not set before, it is set to `.config` by default:

```
KCONFIG_CONFIG    ?= .config
export KCONFIG_CONFIG
```

and the shell that will be used during kernel compilation:

```
CONFIG_SHELL := $(shell if [ -x "$$BASH" ]; then echo $$BASH; \
      else if [ -x /bin/bash ]; then echo /bin/bash; \
      else echo sh; fi ; fi)
```

The next set of variables are related to the compilers used during Linux kernel compilation. We set the host compilers for the `c` and `c++` and the flags to be used with them:

```
HOSTCC       = gcc
HOSTCXX      = g++
HOSTCFLAGS   = -Wall -Wmissing-prototypes -Wstrict-prototypes -O2 -fomit-frame-pointer -std=gnu89
```

```
HOSTCXXFLAGS = -O2
```

Next we get to the `CC` variable that represents compiler too, so why do we need the `HOST*` variables? `CC` is the target compiler that will be used during kernel compilation, but `HOSTCC` will be used during compilation of the set of the `host` programs (we will see it soon). After this we can see the definition of `KBUILD_MODULES` and `KBUILD_BUILTIN` variables that are used to determine what to compile (modules, kernel, or both):

```
KBUILD_MODULES :=
KBUILD_BUILTIN := 1

ifeq ($(MAKECMDGOALS),modules)
  KBUILD_BUILTIN := $(if $(CONFIG_MODVERSIONS),1)
endif
```

Here we can see definition of these variables and the value of `KBUILD_BUILTIN` variable will depend on the `CONFIG_MODVERSIONS` kernel configuration parameter if we pass only `modules` to `make`. The next step is to include the `kbuild` file.

```
include scripts/Kbuild.include
```

The Kbuild or `Kernel Build System` is a special infrastructure to manage building the kernel and its modules. `kbuild` files have the same syntax as makefiles. The scripts/Kbuild.include file provides some generic definitions for the `kbuild` system. After including this `kbuild` file (back in makefile) we can see the definitions of the variables that are related to the different tools used during kernel and module compilation (like linker, compilers, utils from the binutils, etc...):

```
AS        = $(CROSS_COMPILE)as
LD        = $(CROSS_COMPILE)ld
CC        = $(CROSS_COMPILE)gcc
CPP        = $(CC) -E
AR        = $(CROSS_COMPILE)ar
NM        = $(CROSS_COMPILE)nm
STRIP        = $(CROSS_COMPILE)strip
OBJCOPY        = $(CROSS_COMPILE)objcopy
OBJDUMP        = $(CROSS_COMPILE)objdump
AWK        = awk
...
...
...
```

We then define two other variables: `USERINCLUDE` and `LINUXINCLUDE`, which specify paths to header file directories (public for users in the first case and for kernel in the second case):

```
USERINCLUDE    := \
        -I$(srctree)/arch/$(hdr-arch)/include/uapi \
        -Iarch/$(hdr-arch)/include/generated/uapi \
        -I$(srctree)/include/uapi \
        -Iinclude/generated/uapi \
        -include $(srctree)/include/linux/kconfig.h

LINUXINCLUDE    := \
        -I$(srctree)/arch/$(hdr-arch)/include \
        ...
```

And the standard flags for the C compiler:

```
KBUILD_CFLAGS   := -Wall -Wundef -Wstrict-prototypes -Wno-trigraphs \
           -fno-strict-aliasing -fno-common \
           -Werror-implicit-function-declaration \
```

```
           -Wno-format-security \
           -std=gnu89
```

These are not the final compilation flags, as they can be updated in other makefiles (for example kbuilds from `arch/`). After all of these, all variables will be exported to be available in the other makefiles. The `RCS_FIND_IGNORE` and the `RCS_TAR_IGNORE` variables contain files that will be ignored in the version control system:

```
export RCS_FIND_IGNORE := \( -name SCCS -o -name BitKeeper -o -name .svn -o    \
              -name CVS -o -name .pc -o -name .hg -o -name .git \) \
              -prune -o
export RCS_TAR_IGNORE := --exclude SCCS --exclude BitKeeper --exclude .svn \
             --exclude CVS --exclude .pc --exclude .hg --exclude .git
```

With that, we have finished all preparations. The next step is building the `vmlinux` target.

Directly to the kernel build

We have now finished all the preparations, and next step in the main makefile is related to the kernel build. Before this moment, nothing has been printed to the terminal by `make` . But now the first steps of the compilation are started. We need to go to line 598 of the Linux kernel top makefile and we will find the `vmlinux` target there:

```
all: vmlinux
    include arch/$(SRCARCH)/Makefile
```

Don't worry that we have missed many lines in Makefile that are between `export RCS_FIND_IGNORE.....` and `all: vmlinux.....` . This part of the makefile is responsible for the `make *.config` targets and as I wrote in the beginning of this part we will see only building of the kernel in a general way.

The `all:` target is the default when no target is given on the command line. You can see here that we include architecture specific makefile there (in our case it will be arch/x86/Makefile). From this moment we will continue from this makefile. As we can see `all` target depends on the `vmlinux` target that defined a little lower in the top makefile:

```
vmlinux: scripts/link-vmlinux.sh $(vmlinux-deps) FORCE
```

The `vmlinux` is the Linux kernel in a statically linked executable file format. The scripts/link-vmlinux.sh script links and combines different compiled subsystems into vmlinux. The second target is the `vmlinux-deps` that defined as:

```
vmlinux-deps := $(KBUILD_LDS) $(KBUILD_VMLINUX_INIT) $(KBUILD_VMLINUX_MAIN)
```

and consists from the set of the `built-in.o` from each top directory of the Linux kernel. Later, when we will go through all directories in the Linux kernel, the `Kbuild` will compile all the `$(obj-y)` files. It then calls `$(LD) -r` to merge these files into one `built-in.o` file. For this moment we have no `vmlinux-deps` , so the `vmlinux` target will not be executed now. For me `vmlinux-deps` contains following files:

```
arch/x86/kernel/vmlinux.lds arch/x86/kernel/head_64.o
arch/x86/kernel/head64.o    arch/x86/kernel/head.o
init/built-in.o             usr/built-in.o
arch/x86/built-in.o         kernel/built-in.o
mm/built-in.o               fs/built-in.o
ipc/built-in.o              security/built-in.o
crypto/built-in.o           block/built-in.o
lib/lib.a                   arch/x86/lib/lib.a
lib/built-in.o              arch/x86/lib/built-in.o
drivers/built-in.o          sound/built-in.o
firmware/built-in.o         arch/x86/pci/built-in.o
```

```
arch/x86/power/built-in.o   arch/x86/video/built-in.o
net/built-in.o
```

The next target that can be executed is following:

```
$(sort $(vmlinux-deps)): $(vmlinux-dirs) ;
$(vmlinux-dirs): prepare scripts
	$(Q)$(MAKE) $(build)=$@
```

As we can see `vmlinux-dirs` depends on two targets: `prepare` and `scripts` . `prepare` is defined in the top `Makefile` of the Linux kernel and executes three stages of preparations:

```
prepare: prepare0
prepare0: archprepare FORCE
	$(Q)$(MAKE) $(build)=.
archprepare: archheaders archscripts prepare1 scripts_basic

prepare1: prepare2 $(version_h) include/generated/utsrelease.h \
                   include/config/auto.conf
	$(cmd_crmodverdir)
prepare2: prepare3 outputmakefile asm-generic
```

The first `prepare0` expands to the `archprepare` that expands to the `archheaders` and `archscripts` that defined in the `x86_64` specific Makefile. Let's look on it. The `x86_64` specific makefile starts from the definition of the variables that are related to the architecture-specific configs (defconfig, etc...). After this it defines flags for the compiling of the 16-bit code, calculating of the `BITS` variable that can be `32` for `i386` or `64` for the `x86_64` flags for the assembly source code, flags for the linker and many many more (all definitions you can find in the arch/x86/Makefile). The first target is `archheaders` in the makefile generates syscall table:

```
archheaders:
	$(Q)$(MAKE) $(build)=arch/x86/entry/syscalls all
```

And the second target is `archscripts` in this makefile is:

```
archscripts: scripts_basic
	$(Q)$(MAKE) $(build)=arch/x86/tools relocs
```

We can see that it depends on the `scripts_basic` target from the top Makefile. At the first we can see the `scripts_basic` target that executes make for the scripts/basic makefile:

```
scripts_basic:
	$(Q)$(MAKE) $(build)=scripts/basic
```

The `scripts/basic/Makefile` contains targets for compilation of the two host programs: `fixdep` and `bin2` :

```
hostprogs-y    := fixdep
hostprogs-$(CONFIG_BUILD_BIN2C)     += bin2c
always        := $(hostprogs-y)

$(addprefix $(obj)/,$(filter-out fixdep,$(always))): $(obj)/fixdep
```

First program is `fixdep` - optimizes list of dependencies generated by gcc that tells make when to remake a source code file. The second program is `bin2c` , which depends on the value of the `CONFIG_BUILD_BIN2C` kernel configuration option and is a very little C program that allows to convert a binary on stdin to a C include on stdout. You can note here a strange notation: `hostprogs-y` , etc... This notation is used in the all `kbuild` files and you can read more about it in the documentation. In our case `hostprogs-y` tells `kbuild` that there is one host program named `fixdep` that will be built from `fixdep.c` that is located in the same directory where the `Makefile` is. The first output after we execute `make` in our terminal will be result of this `kbuild` file:

```
$ make
  HOSTCC  scripts/basic/fixdep
```

As `script_basic` target was executed, the `archscripts` target will execute `make` for the arch/x86/tools makefile with the `relocs` target:

```
$(Q)$(MAKE) $(build)=arch/x86/tools relocs
```

The `relocs_32.c` and the `relocs_64.c` will be compiled that will contain relocation information and we will see it in the `make` output:

```
  HOSTCC  arch/x86/tools/relocs_32.o
  HOSTCC  arch/x86/tools/relocs_64.o
  HOSTCC  arch/x86/tools/relocs_common.o
  HOSTLD  arch/x86/tools/relocs
```

There is checking of the `version.h` after compiling of the `relocs.c` :

```
$(version_h): $(srctree)/Makefile FORCE
    $(call filechk,version.h)
    $(Q)rm -f $(old_version_h)
```

We can see it in the output:

```
CHK     include/config/kernel.release
```

and the building of the `generic` assembly headers with the `asm-generic` target from the `arch/x86/include/generated/asm` that generated in the top Makefile of the Linux kernel. After the `asm-generic` target the `archprepare` will be done, so the `prepare0` target will be executed. As I wrote above:

```
prepare0: archprepare FORCE
    $(Q)$(MAKE) $(build)=.
```

Note on the `build` . It defined in the scripts/Kbuild.include and looks like this:

```
build := -f $(srctree)/scripts/Makefile.build obj
```

Or in our case it is current source directory - `.` :

```
$(Q)$(MAKE) -f $(srctree)/scripts/Makefile.build obj=.
```

The scripts/Makefile.build tries to find the `Kbuild` file by the given directory via the `obj` parameter, include this `Kbuild` files:

```
include $(kbuild-file)
```

and build targets from it. In our case `.` contains the Kbuild file that generates the `kernel/bounds.s` and the `arch/x86/kernel/asm-offsets.s`. After this the `prepare` target finished to work. The `vmlinux-dirs` also depends on the second target - `scripts` that compiles following programs: `file2alias`, `mk_elfconfig`, `modpost`, etc..... After scripts/host-programs compilation our `vmlinux-dirs` target can be executed. First of all let's try to understand what does `vmlinux-dirs` contain. For my case it contains paths of the following kernel directories:

```
init usr arch/x86 kernel mm fs ipc security crypto block
drivers sound firmware arch/x86/pci arch/x86/power
arch/x86/video net lib arch/x86/lib
```

We can find definition of the `vmlinux-dirs` in the top Makefile of the Linux kernel:

```
vmlinux-dirs    := $(patsubst %/,%,$(filter %/, $(init-y) $(init-m) \
             $(core-y) $(core-m) $(drivers-y) $(drivers-m) \
             $(net-y) $(net-m) $(libs-y) $(libs-m)))

init-y        := init/
drivers-y    := drivers/ sound/ firmware/
net-y        := net/
libs-y        := lib/
...
...
...
```

Here we remove the `/` symbol from the each directory with the help of the `patsubst` and `filter` functions and put it to the `vmlinux-dirs`. So we have list of directories in the `vmlinux-dirs` and the following code:

```
$(vmlinux-dirs): prepare scripts
    $(Q)$(MAKE) $(build)=$@
```

The `$@` represents `vmlinux-dirs` here that means that it will go recursively over all directories from the `vmlinux-dirs` and its internal directories (depens on configuration) and will execute `make` in there. We can see it in the output:

```
  CC      init/main.o
  CHK     include/generated/compile.h
  CC      init/version.o
  CC      init/do_mounts.o
  ...
  CC      arch/x86/crypto/glue_helper.o
  AS      arch/x86/crypto/aes-x86_64-asm_64.o
  CC      arch/x86/crypto/aes_glue.o
  ...
  AS      arch/x86/entry/entry_64.o
  AS      arch/x86/entry/thunk_64.o
  CC      arch/x86/entry/syscall_64.o
```

Source code in each directory will be compiled and linked to the `built-in.o`:

```
$ find . -name built-in.o
./arch/x86/crypto/built-in.o
./arch/x86/crypto/sha-mb/built-in.o
./arch/x86/net/built-in.o
./init/built-in.o
./usr/built-in.o
...
...
```

Ok, all buint-in.o(s) built, now we can back to the `vmlinux` target. As you remember, the `vmlinux` target is in the top Makefile of the Linux kernel. Before the linking of the `vmlinux` it builds samples, Documentation, etc... but I will not describe it here as I wrote in the beginning of this part.

```
vmlinux: scripts/link-vmlinux.sh $(vmlinux-deps) FORCE
    ...
    ...
    +$(call if_changed,link-vmlinux)
```

As you can see main purpose of it is a call of the scripts/link-vmlinux.sh script is linking of the all `built-in.o` (s) to the one statically linked executable and creation of the System.map. In the end we will see following output:

```
  LINK    vmlinux
  LD      vmlinux.o
  MODPOST vmlinux.o
  GEN     .version
  CHK     include/generated/compile.h
  UPD     include/generated/compile.h
  CC      init/version.o
  LD      init/built-in.o
  KSYM    .tmp_kallsyms1.o
  KSYM    .tmp_kallsyms2.o
  LD      vmlinux
  SORTEX  vmlinux
  SYSMAP  System.map
```

and `vmlinux` and `System.map` in the root of the Linux kernel source tree:

```
$ ls vmlinux System.map
System.map  vmlinux
```

That's all, `vmlinux` is ready. The next step is creation of the bzImage.

Building bzImage

The `bzImage` file is the compressed Linux kernel image. We can get it by executing `make bzImage` after `vmlinux` is built. That, or we can just execute `make` without any argument and we will get `bzImage` anyway because it is default image:

```
all: bzImage
```

in the arch/x86/kernel/Makefile. Let's look on this target, it will help us to understand how this image builds. As I already said the `bzImage` target defined in the arch/x86/kernel/Makefile and looks like this:

```
bzImage: vmlinux
    $(Q)$(MAKE) $(build)=$(boot) $(KBUILD_IMAGE)
    $(Q)mkdir -p $(objtree)/arch/$(UTS_MACHINE)/boot
    $(Q)ln -fsn ../../x86/boot/bzImage $(objtree)/arch/$(UTS_MACHINE)/boot/$@
```

We can see here, that first of all called `make` for the boot directory, in our case it is:

```
boot := arch/x86/boot
```

The main goal now is to build the source code in the `arch/x86/boot` and `arch/x86/boot/compressed` directories, build `setup.bin` and `vmlinux.bin`, and build the `bzImage` from them in the end. First target in the arch/x86/boot/Makefile is the `$(obj)/setup.elf`:

```
$(obj)/setup.elf: $(src)/setup.ld $(SETUP_OBJS) FORCE
    $(call if_changed,ld)
```

We already have the `setup.ld` linker script in the `arch/x86/boot` directory and the `SETUP_OBJS` variable that expands to the all source files from the `boot` directory. We can see first output:

```
  AS      arch/x86/boot/bioscall.o
  CC      arch/x86/boot/cmdline.o
  AS      arch/x86/boot/copy.o
  HOSTCC  arch/x86/boot/mkcpustr
  CPUSTR  arch/x86/boot/cpustr.h
  CC      arch/x86/boot/cpu.o
  CC      arch/x86/boot/cpuflags.o
  CC      arch/x86/boot/cpucheck.o
  CC      arch/x86/boot/early_serial_console.o
  CC      arch/x86/boot/edd.o
```

The next source file is arch/x86/boot/header.S, but we can't build it now because this target depends on the following two header files:

```
$(obj)/header.o: $(obj)/voffset.h $(obj)/zoffset.h
```

The first is `voffset.h` generated by the `sed` script that gets two addresses from the `vmlinux` with the `nm` util:

```
#define VO__end 0xffffffff82ab0000
#define VO__text 0xffffffff81000000
```

They are the start and the end of the kernel. The second is `zoffset.h` depens on the `vmlinux` target from the arch/x86/boot/compressed/Makefile:

```
$(obj)/zoffset.h: $(obj)/compressed/vmlinux FORCE
    $(call if_changed,zoffset)
```

The `$(obj)/compressed/vmlinux` target depends on the `vmlinux-objs-y` that compiles source code files from the arch/x86/boot/compressed directory and generates `vmlinux.bin`, `vmlinux.bin.bz2`, and compiles program - `mkpiggy`. We can see this in the output:

```
  LDS     arch/x86/boot/compressed/vmlinux.lds
  AS      arch/x86/boot/compressed/head_64.o
  CC      arch/x86/boot/compressed/misc.o
  CC      arch/x86/boot/compressed/string.o
  CC      arch/x86/boot/compressed/cmdline.o
  OBJCOPY arch/x86/boot/compressed/vmlinux.bin
  BZIP2   arch/x86/boot/compressed/vmlinux.bin.bz2
  HOSTCC  arch/x86/boot/compressed/mkpiggy
```

Where `vmlinux.bin` is the `vmlinux` file with debugging information and comments stripped and the `vmlinux.bin.bz2` compressed `vmlinux.bin.all` + `u32` size of `vmlinux.bin.all`. The `vmlinux.bin.all` is `vmlinux.bin + vmlinux.relocs`, where `vmlinux.relocs` is the `vmlinux` that was handled by the `relocs` program (see above). As we got these files, the `piggy.S` assembly files will be generated with the `mkpiggy` program and compiled:

```
  MKPIGGY arch/x86/boot/compressed/piggy.S
  AS      arch/x86/boot/compressed/piggy.o
```

This assembly files will contain the computed offset from the compressed kernel. After this we can see that `zoffset` generated:

```
  ZOFFSET arch/x86/boot/zoffset.h
```

As the `zoffset.h` and the `voffset.h` are generated, compilation of the source code files from the arch/x86/boot can be continued:

```
  AS      arch/x86/boot/header.o
  CC      arch/x86/boot/main.o
  CC      arch/x86/boot/mca.o
  CC      arch/x86/boot/memory.o
  CC      arch/x86/boot/pm.o
  AS      arch/x86/boot/pmjump.o
  CC      arch/x86/boot/printf.o
  CC      arch/x86/boot/regs.o
  CC      arch/x86/boot/string.o
  CC      arch/x86/boot/tty.o
  CC      arch/x86/boot/video.o
  CC      arch/x86/boot/video-mode.o
  CC      arch/x86/boot/video-vga.o
  CC      arch/x86/boot/video-vesa.o
  CC      arch/x86/boot/video-bios.o
```

As all source code files will be compiled, they will be linked to the `setup.elf` :

```
  LD      arch/x86/boot/setup.elf
```

or:

```
ld -m elf_x86_64   -T arch/x86/boot/setup.ld arch/x86/boot/a20.o arch/x86/boot/bioscall.o arch/x86/boot/cmdline
.o arch/x86/boot/copy.o arch/x86/boot/cpu.o arch/x86/boot/cpuflags.o arch/x86/boot/cpucheck.o arch/x86/boot/ear
ly_serial_console.o arch/x86/boot/edd.o arch/x86/boot/header.o arch/x86/boot/main.o arch/x86/boot/mca.o arch/x8
6/boot/memory.o arch/x86/boot/pm.o arch/x86/boot/pmjump.o arch/x86/boot/printf.o arch/x86/boot/regs.o arch/x86/
boot/string.o arch/x86/boot/tty.o arch/x86/boot/video.o arch/x86/boot/video-mode.o arch/x86/boot/version.o arch
/x86/boot/video-vga.o arch/x86/boot/video-vesa.o arch/x86/boot/video-bios.o -o arch/x86/boot/setup.elf
```

The last two things is the creation of the `setup.bin` that will contain compiled code from the `arch/x86/boot/*` directory:

```
objcopy  -O binary arch/x86/boot/setup.elf arch/x86/boot/setup.bin
```

and the creation of the `vmlinux.bin` from the `vmlinux` :

```
objcopy  -O binary -R .note -R .comment -S arch/x86/boot/compressed/vmlinux arch/x86/boot/vmlinux.bin
```

In the end we compile host program: arch/x86/boot/tools/build.c that will create our `bzImage` from the `setup.bin` and the `vmlinux.bin` :

```
arch/x86/boot/tools/build arch/x86/boot/setup.bin arch/x86/boot/vmlinux.bin arch/x86/boot/zoffset.h arch/x86/bo
ot/bzImage
```

Actually the `bzImage` is the concatenated `setup.bin` and the `vmlinux.bin` . In the end we will see the output which is familiar to all who once built the Linux kernel from source:

```
Setup is 16268 bytes (padded to 16384 bytes).
System is 4704 kB
CRC 94a88f9a
Kernel: arch/x86/boot/bzImage is ready  (#5)
```

That's all.

Conclusion

It is the end of this part and here we saw all steps from the execution of the `make` command to the generation of the `bzImage`. I know, the Linux kernel makefiles and process of the Linux kernel building may seem confusing at first glance, but it is not so hard. Hope this part will help you understand the process of building the Linux kernel.

Links

- GNU make util
- Linux kernel top Makefile
- cross-compilation
- Ctags
- sparse
- bzImage
- uname
- shell
- Kbuild
- binutils
- gcc
- Documentation
- System.map
- Relocation

Introduction

During the writing of the linux-insides book I have received many emails with questions related to the linker script and linker-related subjects. So I've decided to write this to cover some aspects of the linker and the linking of object files.

If we open the `Linker` page on Wikipedia, we will see following definition:

> In computer science, a linker or link editor is a computer program that takes one or more object files generated by a compiler and combines them into a single executable file, library file, or another object file.

If you've written at least one program on C in your life, you will have seen files with the `*.o` extension. These files are object files. Object files are blocks of machine code and data with placeholder addresses that reference data and functions in other object files or libraries, as well as a list of its own functions and data. The main purpose of the linker is collect/handle the code and data of each object file, turning it into the final executable file or library. In this post we will try to go through all aspects of this process. Let's start.

Linking process

Let's create a simple project with the following structure:

```
*-linkers
*--main.c
*--lib.c
*--lib.h
```

Our `main.c` source code file contains:

```
#include <stdio.h>

#include "lib.h"

int main(int argc, char **argv) {
    printf("factorial of 5 is: %d\n", factorial(5));
    return 0;
}
```

The `lib.c` file contains:

```
int factorial(int base) {
    int res,i = 1;

    if (base == 0) {
        return 1;
    }

    while (i <= base) {
        res *= i;
        i++;
    }

    return res;
}
```

And the `lib.h` file contains:

```
#ifndef LIB_H
```

```
#define LIB_H

int factorial(int base);

#endif
```

Now let's compile only the `main.c` source code file with:

```
$ gcc -c main.c
```

If we look inside the outputted object file with the `nm` util, we will see the following output:

```
$ nm -A main.o
main.o:                 U factorial
main.o:0000000000000000 T main
main.o:                 U printf
```

The `nm` util allows us to see the list of symbols from the given object file. It consists of three columns: the first is the name of the given object file and the address of any resolved symbols. The second column contains a character that represents the status of the given symbol. In this case the `U` means `undefined` and the `T` denotes that the symbols are placed in the `.text` section of the object. The `nm` utility shows us here that we have three symbols in the `main.c` source code file:

- `factorial` - the factorial function defined in the `lib.c` source code file. It is marked as `undefined` here because we compiled only the `main.c` source code file, and it does not know anything about code from the `lib.c` file for now;
- `main` - the main function;
- `printf` - the function from the glibc library. `main.c` does not know anything about it for now either.

What can we understand from the output of `nm` so far? The `main.o` object file contains the local symbol `main` at address `0000000000000000` (it will be filled with correct address after is is linked), and two unresolved symbols. We can see all of this information in the disassembly output of the `main.o` object file:

```
$ objdump -S main.o

main.o:     file format elf64-x86-64
Disassembly of section .text:

0000000000000000 <main>:
   0:	55                   	push   %rbp
   1:	48 89 e5             	mov    %rsp,%rbp
   4:	48 83 ec 10          	sub    $0x10,%rsp
   8:	89 7d fc             	mov    %edi,-0x4(%rbp)
   b:	48 89 75 f0          	mov    %rsi,-0x10(%rbp)
   f:	bf 05 00 00 00       	mov    $0x5,%edi
  14:	e8 00 00 00 00       	callq  19 <main+0x19>
  19:	89 c6                	mov    %eax,%esi
  1b:	bf 00 00 00 00       	mov    $0x0,%edi
  20:	b8 00 00 00 00       	mov    $0x0,%eax
  25:	e8 00 00 00 00       	callq  2a <main+0x2a>
  2a:	b8 00 00 00 00       	mov    $0x0,%eax
  2f:	c9                   	leaveq
  30:	c3                   	retq
```

Here we are interested only in the two `callq` operations. The two `callq` operations contain `linker stubs`, or the function name and offset from it to the next instruction. These stubs will be updated to the real addresses of the functions. We can see these functions' names with in the following `objdump` output:

```
$ objdump -S -r main.o

...
```

```
  14:    e8 00 00 00 00           callq  19 <main+0x19>
  15: R_X86_64_PC32                   factorial-0x4
  19:    89 c6                    mov    %eax,%esi
...
  25:    e8 00 00 00 00           callq  2a <main+0x2a>
  26:   R_X86_64_PC32                     printf-0x4
  2a:    b8 00 00 00 00           mov    $0x0,%eax
...
```

The `-r` or `--reloc` flags of the `objdump` util print the `relocation` entries of the file. Now let's look in more detail at the relocation process.

Relocation

Relocation is the process of connecting symbolic references with symbolic definitions. Let's look at the previous snippet from the `objdump` output:

```
  14:    e8 00 00 00 00           callq  19 <main+0x19>
  15:   R_X86_64_PC32                     factorial-0x4
  19:    89 c6                    mov    %eax,%esi
```

Note the `e8 00 00 00 00` on the first line. The `e8` is the opcode of the `call`, and the remainder of the line is a relative offset. So the `e8 00 00 00 00` contains a one-byte operation code followed by a four-byte address. Note that the `00 00 00 00` is 4-bytes. Why only 4-bytes if an address can be 8-bytes in a `x86_64` (64-bit) machine? Actually we compiled the `main.c` source code file with the `-mcmodel=small`! From the `gcc` man page:

```
-mcmodel=small

Generate code for the small code model: the program and its symbols must be linked in the lower 2 GB of the add
ress space. Pointers are 64 bits. Programs can be statically or dynamically linked. This is the default code mo
del.
```

Of course we didn't pass this option to the `gcc` when we compiled the `main.c`, but it is the default. We know that our program will be linked in the lower 2 GB of the address space from the `gcc` manual extract above. Four bytes is therefore enough for this. So we have opcode of the `call` instruction and an unknown address. When we compile `main.c` with all its dependencies to an executable file, and then look at the factorial call we see:

```
$ gcc main.c lib.c -o factorial | objdump -S factorial | grep factorial

factorial:     file format elf64-x86-64
...
...
0000000000400506 <main>:
    40051a:    e8 18 00 00 00           callq  400537 <factorial>
...
...
0000000000400537 <factorial>:
    400550:    75 07                    jne    400559 <factorial+0x22>
    400557:    eb 1b                    jmp    400574 <factorial+0x3d>
    400559:    eb 0e                    jmp    400569 <factorial+0x32>
    40056f:    7e ea                    jle    40055b <factorial+0x24>
...
...
```

As we can see in the previous output, the address of the `main` function is `0x0000000000400506`. Why it does not start from `0x0`? You may already know that standard C programs are linked with the `glibc` C standard library (assuming the `-nostdlib` was not passed to the `gcc`). The compiled code for a program includes constructor functions to initialize data in the program when the program is started. These functions need to be called before the program is started, or in another words before the `main` function is called. To make the initialization and termination functions work, the compiler must output something in the assembler code to cause those functions to be called at the appropriate time. Execution of this program will start from the code placed in the special `.init` section. We can see this in the beginning of the objdump output:

```
objdump -S factorial | less

factorial:     file format elf64-x86-64

Disassembly of section .init:

00000000004003a8 <_init>:
  4003a8:       48 83 ec 08             sub    $0x8,%rsp
  4003ac:       48 8b 05 a5 05 20 00    mov    0x2005a5(%rip),%rax        # 600958 <_DYNAMIC+0x1d0>
```

Not that it starts at the `0x00000000004003a8` address relative to the `glibc` code. We can check it also in the ELF output by running `readelf` :

```
$ readelf -d factorial | grep \(INIT\)
 0x000000000000000c (INIT)               0x4003a8
```

So, the address of the `main` function is `0000000000400506` and is offset from the `.init` section. As we can see from the output, the address of the `factorial` function is `0x0000000000400537` and binary code for the call of the `factorial` function now is `e8 18 00 00 00` . We already know that `e8` is opcode for the `call` instruction, the next `18 00 00 00` (note that address represented as little endian for `x86_64` , so it is `00 00 00 18`) is the offset from the `callq` to the `factorial` function:

```
>>> hex(0x40051a + 0x18 + 0x5) == hex(0x400537)
True
```

So we add `0x18` and `0x5` to the address of the `call` instruction. The offset is measured from the address of the following instruction. Our call instruction is 5-bytes long (`e8 18 00 00 00`) and the `0x18` is the offset of the call after the `factorial` function. A compiler generally creates each object file with the program addresses starting at zero. But if a program is created from multiple object files, these will overlap.

What we have seen in this section is the `relocation` process. This process assigns load addresses to the various parts of the program, adjusting the code and data in the program to reflect the assigned addresses.

Ok, now that we know a little about linkers and relocation it is time to learn more about linkers by linking our object files.

GNU linker

As you can understand from the title, I will use GNU linker or just `ld` in this post. Of course we can use `gcc` to link our `factorial` project:

```
$ gcc main.c lib.o -o factorial
```

and after it we will get executable file - `factorial` as a result:

```
./factorial
factorial of 5 is: 120
```

But `gcc` does not link object files. Instead it uses `collect2` which is just wrapper for the `GNU ld` linker:

```
~$ /usr/lib/gcc/x86_64-linux-gnu/4.9/collect2 --version
```

```
collect2 version 4.9.3
/usr/bin/ld --version
GNU ld (GNU Binutils for Debian) 2.25
...
...
...
```

Ok, we can use gcc and it will produce executable file of our program for us. But let's look how to use `GNU ld` linker for the same purpose. First of all let's try to link these object files with the following example:

```
ld main.o lib.o -o factorial
```

Try to do it and you will get following error:

```
$ ld main.o lib.o -o factorial
ld: warning: cannot find entry symbol _start; defaulting to 00000000004000b0
main.o: In function `main':
main.c:(.text+0x26): undefined reference to `printf'
```

Here we can see two problems:

* Linker can't find `_start` symbol;
* Linker does not know anything about `printf` function.

First of all let's try to understand what is this `_start` entry symbol that appears to be required for our program to run? When I started to learn programming I learned that the `main` function is the entry point of the program. I think you learned this too :) But it actually isn't the entry point, it's `_start` instead. The `_start` symbol is defined in the `crt1.o` object file. We can find it with the following command:

```
$ objdump -S /usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o

/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o:     file format elf64-x86-64


Disassembly of section .text:

0000000000000000 <_start>:
   0:    31 ed                    xor    %ebp,%ebp
   2:    49 89 d1                 mov    %rdx,%r9
   ...
   ...
   ...
```

We pass this object file to the `ld` command as its first argument (see above). Now let's try to link it and will look on result:

```
ld /usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o \
main.o lib.o -o factorial

/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o: In function `_start':
/tmp/buildd/glibc-2.19/csu/../sysdeps/x86_64/start.S:115: undefined reference to `__libc_csu_fini'
/tmp/buildd/glibc-2.19/csu/../sysdeps/x86_64/start.S:116: undefined reference to `__libc_csu_init'
/tmp/buildd/glibc-2.19/csu/../sysdeps/x86_64/start.S:122: undefined reference to `__libc_start_main'
main.o: In function `main':
main.c:(.text+0x26): undefined reference to `printf'
```

Unfortunately we will see even more errors. We can see here old error about undefined `printf` and yet another three undefined references:

* `__libc_csu_fini`

* `__libc_csu_init`
* `__libc_start_main`

The `_start` symbol is defined in the sysdeps/x86_64/start.S assembly file in the `glibc` source code. We can find following assembly code lines there:

```
mov $__libc_csu_fini, %R8_LP
mov $__libc_csu_init, %RCX_LP
...
call __libc_start_main
```

Here we pass address of the entry point to the `.init` and `.fini` section that contain code that starts to execute when the program is ran and the code that executes when program terminates. And in the end we see the call of the `main` function from our program. These three symbols are defined in the csu/elf-init.c source code file. The following two object files:

* `crtn.o` ;
* `crti.o` .

define the function prologs/epilogs for the .init and .fini sections (with the `_init` and `_fini` symbols respectively).

The `crtn.o` object file contains these `.init` and `.fini` sections:

```
$ objdump -S /usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crtn.o

0000000000000000 <.init>:
   0:    48 83 c4 08             add    $0x8,%rsp
   4:    c3                      retq

Disassembly of section .fini:

0000000000000000 <.fini>:
   0:    48 83 c4 08             add    $0x8,%rsp
   4:    c3                      retq
```

And the `crti.o` object file contains the `_init` and `_fini` symbols. Let's try to link again with these two object files:

```
$ ld \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crti.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crtn.o main.o lib.o \
-o factorial
```

And anyway we will get the same errors. Now we need to pass `-lc` option to the `ld` . This option will search for the standard library in the paths present in the `$LD_LIBRARY_PATH` environment variable. Let's try to link again wit the `-lc` option:

```
$ ld \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crti.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crtn.o main.o lib.o -lc \
-o factorial
```

Finally we get an executable file, but if we try to run it, we will get strange results:

```
$ ./factorial
bash: ./factorial: No such file or directory
```

What's the problem here? Let's look on the executable file with the readelf util:

```
$ readelf -l factorial

Elf file type is EXEC (Executable file)
Entry point 0x4003c0
There are 7 program headers, starting at offset 64

Program Headers:
  Type           Offset             VirtAddr           PhysAddr
                 FileSiz            MemSiz              Flags  Align
  PHDR           0x0000000000000040 0x0000000000400040 0x0000000000400040
                 0x0000000000000188 0x0000000000000188  R E    8
  INTERP         0x00000000000001c8 0x00000000004001c8 0x00000000004001c8
                 0x000000000000001c 0x000000000000001c  R      1
      [Requesting program interpreter: /lib64/ld-linux-x86-64.so.2]
  LOAD           0x0000000000000000 0x0000000000400000 0x0000000000400000
                 0x0000000000000610 0x0000000000000610  R E    200000
  LOAD           0x0000000000000610 0x0000000000600610 0x0000000000600610
                 0x00000000000001cc 0x00000000000001cc  RW     200000
  DYNAMIC        0x0000000000000610 0x0000000000600610 0x0000000000600610
                 0x0000000000000190 0x0000000000000190  RW     8
  NOTE           0x00000000000001e4 0x00000000004001e4 0x00000000004001e4
                 0x0000000000000020 0x0000000000000020  R      4
  GNU_STACK      0x0000000000000000 0x0000000000000000 0x0000000000000000
                 0x0000000000000000 0x0000000000000000  RW     10

 Section to Segment mapping:
  Segment Sections...
   00
   01     .interp
   02     .interp .note.ABI-tag .hash .dynsym .dynstr .gnu.version .gnu.version_r .rela.dyn .rela.plt .init .pl
t .text .fini .rodata .eh_frame
   03     .dynamic .got .got.plt .data
   04     .dynamic
   05     .note.ABI-tag
   06
```

Note on the strange line:

```
  INTERP         0x00000000000001c8 0x00000000004001c8 0x00000000004001c8
                 0x000000000000001c 0x000000000000001c  R      1
      [Requesting program interpreter: /lib64/ld-linux-x86-64.so.2]
```

The `.interp` section in the `elf` file holds the path name of a program interpreter or in another words the `.interp` section simply contains an `ascii` string that is the name of the dynamic linker. The dynamic linker is the part of Linux that loads and links shared libraries needed by an executable when it is executed, by copying the content of libraries from disk to RAM. As we can see in the output of the `readelf` command it is placed in the `/lib64/ld-linux-x86-64.so.2` file for the `x86_64` architecture. Now let's add the `-dynamic-linker` option with the path of `ld-linux-x86-64.so.2` to the `ld` call and will see the following results:

```
$ gcc -c main.c lib.c

$ ld \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crti.o \
/usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crtn.o main.o lib.o \
-dynamic-linker /lib64/ld-linux-x86-64.so.2 \
-lc -o factorial
```

Now we can run it as normal executable file:

```
$ ./factorial
```

```
factorial of 5 is: 120
```

It works! With the first line we compile the `main.c` and the `lib.c` source code files to object files. We will get the `main.o` and the `lib.o` after execution of the `gcc` :

```
$ file lib.o main.o
lib.o:  ELF 64-bit LSB relocatable, x86-64, version 1 (SYSV), not stripped
main.o: ELF 64-bit LSB relocatable, x86-64, version 1 (SYSV), not stripped
```

and after this we link object files of our program with the needed system object files and libraries. We just saw a simple example of how to compile and link a C program with the `gcc` compiler and `GNU ld` linker. In this example we have used a couple command line options of the `GNU linker` , but it supports much more command line options than `-o` , `-dynamic-linker` , etc... Moreover `GNU ld` has its own language that allows to control the linking process. In the next two paragraphs we will look into it.

Useful command line options of the GNU linker

As I already wrote and as you can see in the manual of the `GNU linker` , it has big set of the command line options. We've seen a couple of options in this post: `-o <output>` - that tells `ld` to produce an output file called `output` as the result of linking, `-l<name>` that adds the archive or object file specified by the name, `-dynamic-linker` that specifies the name of the dynamic linker. Of course `ld` supports much more command line options, let's look at some of them.

The first useful command line option is `@file` . In this case the `file` specifies filename where command line options will be read. For example we can create file with the name `linker.ld` , put there our command line arguments from the previous example and execute it with:

```
$ ld @linker.ld
```

The next command line option is `-b` or `--format` . This command line option specifies format of the input object files `ELF` , `DJGPP/COFF` and etc. There is a command line option for the same purpose but for the output file: `--oformat=output-format` .

The next command line option is `--defsym` . Full format of this command line option is the `--defsym=symbol=expression` . It allows to create global symbol in the output file containing the absolute address given by expression. We can find following case where this command line option can be useful: in the Linux kernel source code and more precisely in the Makefile that is related to the kernel decompression for the ARM architecture - arch/arm/boot/compressed/Makefile, we can find following definition:

```
LDFLAGS_vmlinux = --defsym _kernel_bss_size=$(KBSS_SZ)
```

As we already know, it defines the `_kernel_bss_size` symbol with the size of the `.bss` section in the output file. This symbol will be used in the first assembly file that will be executed during kernel decompressing:

```
ldr r5, =_kernel_bss_size
```

The next command line options is the `-shared` that allows us to create shared library. The `-M` or `-map <filename>` command line option prints the linking map with the information about symbols. In our case:

```
$ ld -M @linker.ld
...
...
...
.text           0x00000000004003c0      0x112
 *(.text.unlikely .text.*_unlikely .text.unlikely.*)
 *(.text.exit .text.exit.*)
 *(.text.startup .text.startup.*)
```

```
  *(.text.hot .text.hot.*)
  *(.text .stub .text.* .gnu.linkonce.t.*)
  .text          0x00000000004003c0       0x2a /usr/lib/gcc/x86_64-linux-gnu/4.9/../../../x86_64-linux-gnu/crt1.
o
...
...
...
  .text          0x00000000004003ea       0x31 main.o
                 0x00000000004003ea                main
  .text          0x000000000040041b       0x3f lib.o
                 0x000000000040041b                factorial
```

Of course the `GNU linker` support standard command line options: `--help` and `--version` that print common help of the usage of the `ld` and its version. That's all about command line options of the `GNU linker`. Of course it is not the full set of command line options supported by the `ld` util. You can find the complete documentation of the `ld` util in the manual.

Control Language linker

As I wrote previously, `ld` has support for its own language. It accepts Linker Command Language files written in a superset of AT&T's Link Editor Command Language syntax, to provide explicit and total control over the linking process. Let's look on its details.

With the linker language we can control:

- input files;
- output files;
- file formats
- addresses of sections;
- etc...

Commands written in the linker control language are usually placed in a file called linker script. We can pass it to `ld` with the `-T` command line option. The main command in a linker script is the `SECTIONS` command. Each linker script must contain this command and it determines the `map` of the output file. The special variable `.` contains current position of the output. Let's write a simple assembly program and we will look at how we can use a linker script to control linking of this program. We will take a hello world program for this example:

```
.data
        msg:    .ascii  "hello, world!\n"

.text

.global _start

_start:
        mov    $1,%rax
        mov    $1,%rdi
        mov    $msg,%rsi
        mov    $14,%rdx
        syscall

        mov    $60,%rax
        mov    $0,%rdi
        syscall
```

We can compile and link it with the following commands:

```
$ as -o hello.o hello.asm
$ ld -o hello hello.o
```

Our program consists from two sections: `.text` contains code of the program and `.data` contains initialized variables. Let's write simple linker script and try to link our `hello.asm` assembly file with it. Our script is:

```
/*
 * Linker script for the factorial
 */
OUTPUT(hello)
OUTPUT_FORMAT("elf64-x86-64")
INPUT(hello.o)

SECTIONS
{
    . = 0x200000;
    .text : {
          *(.text)
    }

    . = 0x400000;
    .data : {
          *(.data)
    }
}
```

On the first three lines you can see a comment written in `C` style. After it the `OUTPUT` and the `OUTPUT_FORMAT` commands specify the name of our executable file and its format. The next command, `INPUT`, specifies the input file to the `ld` linker. Then, we can see the main `SECTIONS` command, which, as I already wrote, must be present in every linker script. The `SECTIONS` command represents the set and order of the sections which will be in the output file. At the beginning of the `SECTIONS` command we can see following line `. = 0x200000`. I already wrote above that `.` command points to the current position of the output. This line says that the code should be loaded at address `0x200000` and the line `. = 0x400000` says that data section should be loaded at address `0x400000`. The second line after the `. = 0x200000` defines `.text` as an output section. We can see `*(.text)` expression inside it. The `*` symbol is wildcard that matches any file name. In other words, the `*(.text)` expression says all `.text` input sections in all input files. We can rewrite it as `hello.o(.text)` for our example. After the following location counter `. = 0x400000`, we can see definition of the data section.

We can compile and link it with the following command:

```
$ as -o hello.o hello.S && ld -T linker.script && ./hello
hello, world!
```

If we look inside it with the `objdump` util, we can see that `.text` section starts from the address `0x200000` and the `.data` sections starts from the address `0x400000`:

```
$ objdump -D hello

Disassembly of section .text:

0000000000200000 <_start>:
  200000:    48 c7 c0 01 00 00 00     mov    $0x1,%rax
  ...

Disassembly of section .data:

0000000000400000 <msg>:
  400000:    68 65 6c 6c 6f           pushq  $0x6f6c6c65
  ...
```

Apart from the commands we have already seen, there are a few others. The first is the `ASSERT(exp, message)` that ensures that given expression is not zero. If it is zero, then exit the linker with an error code and print the given error message. If you've read about Linux kernel booting process in the linux-insides book, you may know that the setup header of the Linux kernel has offset

`0x1f1` . In the linker script of the Linux kernel we can find a check for this:

```
. = ASSERT(hdr == 0x1f1, "The setup header has the wrong offset!");
```

The `INCLUDE filename` command allows to include external linker script symbols in the current one. In a linker script we can assign a value to a symbol. `ld` supports a couple of assignment operators:

- symbol = expression ;
- symbol += expression ;
- symbol -= expression ;
- symbol *= expression ;
- symbol /= expression ;
- symbol <<= expression ;
- symbol >>= expression ;
- symbol &= expression ;
- symbol |= expression ;

As you can note all operators are C assignment operators. For example we can use it in our linker script as:

```
START_ADDRESS = 0x200000;
DATA_OFFSET   = 0x200000;

SECTIONS
{
    . = START_ADDRESS;
    .text : {
          *(.text)
    }

    . = START_ADDRESS + DATA_OFFSET;
    .data : {
          *(.data)
    }
}
```

As you already may noted the syntax for expressions in the linker script language is identical to that of C expressions. Besides this the control language of the linking supports following builtin functions:

- `ABSOLUTE` - returns absolute value of the given expression;
- `ADDR` - takes the section and returns its address;
- `ALIGN` - returns the value of the location counter (`.` operator) that aligned by the boundary of the next expression after the given expression;
- `DEFINED` - returns `1` if the given symbol placed in the global symbol table and `0` in other way;
- `MAX` and `MIN` - return maximum and minimum of the two given expressions;
- `NEXT` - returns the next unallocated address that is a multiple of the give expression;
- `SIZEOF` - returns the size in bytes of the given named section.

That's all.

Conclusion

This is the end of the post about linkers. We learned many things about linkers in this post, such as what is a linker and why it is needed, how to use it, etc..

If you have any questions or suggestions, write me an email or ping me on twitter.

Please note that English is not my first language, and I am really sorry for any inconvenience. If you find any mistakes please let me know via email or send a PR.

Links

- Book about Linux kernel insides
- linker
- object files
- glibc
- opcode
- ELF
- GNU linker
- My posts about assembly programming for x86_64
- readelf

Program startup process in userspace

Introduction

Despite the linux-insides described mostly Linux kernel related stuff, I have decided to write this one part which mostly related to userspace.

There is already fourth part of System calls chapter which describes what does the Linux kernel do when we want to start a program. In this part I want to explore what happens when we run a program on a Linux machine from userspace perspective.

I don't know how about you, but in my university I learn that a `C` program starts executing from the function which is called `main`. And that's partly true. Whenever we are starting to write new program, we start our program from the following lines of code:

```
int main(int argc, char *argv[]) {
    // Entry point is here
}
```

But if you are interested in low-level programming, you may already know that the `main` function isn't the actual entry point of a program. You will believe it's true after you look at this simple program in debugger:

```
int main(int argc, char *argv[]) {
    return 0;
}
```

Let's compile this and run in gdb:

```
$ gcc -ggdb program.c -o program
$ gdb ./program
The target architecture is assumed to be i386:x86-64:intel
Reading symbols from ./program...done.
```

Let's execute gdb `info` subcommand with `files` argument. The `info files` prints information about debugging targets and memory spaces occupied by different sections.

```
(gdb) info files
Symbols from "/home/alex/program".
Local exec file:
    `/home/alex/program', file type elf64-x86-64.
    Entry point: 0x400430
    0x0000000000400238 - 0x0000000000400254 is .interp
    0x0000000000400254 - 0x0000000000400274 is .note.ABI-tag
    0x0000000000400274 - 0x0000000000400298 is .note.gnu.build-id
    0x0000000000400298 - 0x00000000004002b4 is .gnu.hash
    0x00000000004002b8 - 0x0000000000400318 is .dynsym
    0x0000000000400318 - 0x0000000000400357 is .dynstr
    0x0000000000400358 - 0x0000000000400360 is .gnu.version
    0x0000000000400360 - 0x0000000000400380 is .gnu.version_r
    0x0000000000400380 - 0x0000000000400398 is .rela.dyn
    0x0000000000400398 - 0x00000000004003c8 is .rela.plt
    0x00000000004003c8 - 0x00000000004003e2 is .init
    0x00000000004003f0 - 0x0000000000400420 is .plt
    0x0000000000400420 - 0x0000000000400428 is .plt.got
    0x0000000000400430 - 0x00000000004005e2 is .text
    0x00000000004005e4 - 0x00000000004005ed is .fini
    0x00000000004005f0 - 0x0000000000400610 is .rodata
```

```
    0x0000000000400610 - 0x0000000000400644 is .eh_frame_hdr
    0x0000000000400648 - 0x000000000040073c is .eh_frame
    0x0000000000600e10 - 0x0000000000600e18 is .init_array
    0x0000000000600e18 - 0x0000000000600e20 is .fini_array
    0x0000000000600e20 - 0x0000000000600e28 is .jcr
    0x0000000000600e28 - 0x0000000000600ff8 is .dynamic
    0x0000000000600ff8 - 0x0000000000601000 is .got
    0x0000000000601000 - 0x0000000000601028 is .got.plt
    0x0000000000601028 - 0x0000000000601034 is .data
    0x0000000000601034 - 0x0000000000601038 is .bss
```

Note on `Entry point: 0x400430` line. Now we know the actual address of entry point of our program. Let's put a breakpoint by this address, run our program and see what happens:

```
(gdb) break *0x400430
Breakpoint 1 at 0x400430
(gdb) run
Starting program: /home/alex/program

Breakpoint 1, 0x0000000000400430 in _start ()
```

Interesting. We don't see execution of the `main` function here, but we have seen that another function is called. This function is `_start` and as our debugger shows us, it is the actual entry point of our program. Where is this function from? Who does call `main` and when is it called? I will try to answer all these questions in the following post.

How the kernel starts a new program

First of all, let's take a look at the following simple `C` program:

```C
// program.c

#include <stdlib.h>
#include <stdio.h>

static int x = 1;

int y = 2;

int main(int argc, char *argv[]) {
    int z = 3;

    printf("x + y + z = %d\n", x + y + z);

    return EXIT_SUCCESS;
}
```

We can be sure that this program works as we expect. Let's compile it:

```
$ gcc -Wall program.c -o sum
```

and run:

```
$ ./sum
x + y + z = 6
```

Ok, everything looks pretty good up to now. You may already know that there is a special family of functions - exec*. As we read in the man page:

> The exec() family of functions replaces the current process image with a new process image.

All the `exec*` functions are simple frontends to the execve system call. If you have read the fourth part of the chapter which describes system calls, you may know that the execve system call is defined in the files/exec.c source code file and looks like:

```
SYSCALL_DEFINE3(execve,
        const char __user *, filename,
        const char __user *const __user *, argv,
        const char __user *const __user *, envp)
{
    return do_execve(getname(filename), argv, envp);
}
```

It takes an executable file name, set of command line arguments, and set of enviroment variables. As you may guess, everything is done by the `do_execve` function. I will not describe the implementation of the `do_execve` function in detail because you can read about this in here. But in short words, the `do_execve` function does many checks like `filename` is valid, limit of launched processes is not exceed in our system and etc. After all of these checks, this function parses our executable file which is represented in ELF format, creates memory descriptor for newly executed executable file and fills it with the appropriate values like area for the stack, heap and etc. When the setup of new binary image is done, the `start_thread` function will set up one new process. This function is architecture-specific and for the x86_64 architecture, its definition will be located in the arch/x86/kernel/process_64.c source code file.

The `start_thread` function sets new value to segment registers and program execution address. From this point, our new process is ready to start. Once the context switch will be done, control will be returned to userspace with new values of registers and the new executable will be started to execute.

That's all from the kernel side. The Linux kernel prepares the binary image for execution and its execution starts right after the context switch and returns controll to userspace when it is finished. But it does not answer our questions like where does `_start` come from and others. Let's try to answer these questions in the next paragraph.

How does a program start in userspace

In the previous paragraph we saw how an executable file is prepared to run by the Linux kernel. Let's look at the same, but from userspace side. We already know that the entry point of each program is its `_start` function. But where is this function from? It may came from a library. But if you remember correctly we didn't link our program with any libraries during compilation of our program:

```
$ gcc -Wall program.c -o sum
```

You may guess that `_start` comes from the stanard libray and that's true. If you try to compile our program again and pass the `-v` option to gcc which will enable `verbose mode`, you will see a long output. The full output is not interesting for us, let's look at the following steps:

First of all, our program should be compiled with `gcc`:

```
$ gcc -v -ggdb program.c -o sum
...
...
...
/usr/libexec/gcc/x86_64-redhat-linux/6.1.1/cc1 -quiet -v program.c -quiet -dumpbase program.c -mtune=generic -m
arch=x86-64 -auxbase test -ggdb -version -o /tmp/ccvUWZkF.s
...
...
...
```

The `cc1` compiler will compile our `C` source code and an produce assembly named `/tmp/ccvUWZkF.s` file. After this we can see that our assembly file will be compiled into object file with the `GNU as` assembler:

```
$ gcc -v -ggdb program.c -o sum
...
...
...
as -v --64 -o /tmp/cc79wZSU.o /tmp/ccvUWZkF.s
...
...
...
```

In the end our object file will be linked by `collect2` :

```
$ gcc -v -ggdb program.c -o sum
...
...
...
/usr/libexec/gcc/x86_64-redhat-linux/6.1.1/collect2 -plugin /usr/libexec/gcc/x86_64-redhat-linux/6.1.1/liblto_p
lugin.so -plugin-opt=/usr/libexec/gcc/x86_64-redhat-linux/6.1.1/lto-wrapper -plugin-opt=-fresolution=/tmp/ccLEG
Yra.res -plugin-opt=-pass-through=-lgcc -plugin-opt=-pass-through=-lgcc_s -plugin-opt=-pass-through=-lc -plugin
-opt=-pass-through=-lgcc -plugin-opt=-pass-through=-lgcc_s --build-id --no-add-needed --eh-frame-hdr --hash-sty
le=gnu -m elf_x86_64 -dynamic-linker /lib64/ld-linux-x86-64.so.2 -o test /usr/lib/gcc/x86_64-redhat-linux/6.1.1
/../../../../lib64/crt1.o /usr/lib/gcc/x86_64-redhat-linux/6.1.1/../../../../lib64/crti.o /usr/lib/gcc/x86_64-r
edhat-linux/6.1.1/crtbegin.o -L/usr/lib/gcc/x86_64-redhat-linux/6.1.1 -L/usr/lib/gcc/x86_64-redhat-linux/6.1.1/
../../../../lib64 -L/lib/../lib64 -L/usr/lib/../lib64 -L. -L/usr/lib/gcc/x86_64-redhat-linux/6.1.1/../../.. /tm
p/cc79wZSU.o -lgcc --as-needed -lgcc_s --no-as-needed -lc -lgcc --as-needed -lgcc_s --no-as-needed /usr/lib/gcc
/x86_64-redhat-linux/6.1.1/crtend.o /usr/lib/gcc/x86_64-redhat-linux/6.1.1/../../../../lib64/crtn.o
...
...
...
```

Yes, we can see a long set of command line options which are passed to the linker. Let's go from another way. We know that our program depends on `stdlib` :

```
$ ldd program
    linux-vdso.so.1 (0x00007ffc9afd2000)
    libc.so.6 => /lib64/libc.so.6 (0x00007f56b389b000)
    /lib64/ld-linux-x86-64.so.2 (0x0000556198231000)
```

as we use some stuff from there like `printf` and etc. But not only. That's why we will get an error when we pass `-nostdlib` option to the compiler:

```
$ gcc -nostdlib program.c -o program
/usr/bin/ld: warning: cannot find entry symbol _start; defaulting to 000000000040017c
/tmp/cc02msGW.o: In function `main':
/home/alex/program.c:11: undefined reference to `printf'
collect2: error: ld returned 1 exit status
```

Besides other errors, we also see that `_start` symbol is undefined. So now we are sure that the `_start` function comes from standard library. But even if we link it with the standard library, it will not be compiled successfully anyway:

```
$ gcc -nostdlib -lc -ggdb program.c -o program
/usr/bin/ld: warning: cannot find entry symbol _start; defaulting to 0000000000400350
```

Ok, the compiler does not complain about undefined reference of standard library functions anymore as we linked our program with `/usr/lib64/libc.so.6` , but the `_start` symbol isn't resolved yet. Let's return to the verbose output of `gcc` and look at the parameters of `collect2` . The most important thing that we may see is that our program is linked not only with the standard library, but also with some object files. The first object file is: `/lib64/crt1.o` . And if we look inside this object file with `objdump` , we will see the `_start` symbol:

```
$ objdump -d /lib64/crt1.o

/lib64/crt1.o:     file format elf64-x86-64


Disassembly of section .text:

0000000000000000 <_start>:
   0:    31 ed                    xor    %ebp,%ebp
   2:    49 89 d1                 mov    %rdx,%r9
   5:    5e                       pop    %rsi
   6:    48 89 e2                 mov    %rsp,%rdx
   9:    48 83 e4 f0              and    $0xfffffffffffffff0,%rsp
   d:    50                       push   %rax
   e:    54                       push   %rsp
   f:    49 c7 c0 00 00 00 00     mov    $0x0,%r8
  16:    48 c7 c1 00 00 00 00     mov    $0x0,%rcx
  1d:    48 c7 c7 00 00 00 00     mov    $0x0,%rdi
  24:    e8 00 00 00 00           callq  29 <_start+0x29>
  29:    f4                       hlt
```

As `crt1.o` is a shared object file, we see only stubs here instead of real calls. Let's look at the source code of the `_start` function. As this function is architecture specific, implementation for `_start` will be located in the sysdeps/x86_64/start.S assembly file.

The `_start` starts from the clearing of `ebp` register as ABI suggests.

```
xorl %ebp, %ebp
```

And after this we put the address of termination function to the `r9` register:

```
mov %RDX_LP, %R9_LP
```

As described in the ELF specification:

> After the dynamic linker has built the process image and performed the relocations, each shared object gets the opportunity to execute some initialization code. ... Similarly, shared objects may have termination functions, which are executed with the atexit (BA_OS) mechanism after the base process begins its termination sequence.

So we need to put the address of the termination function to the `r9` register as it will be passed to `__libc_start_main` in future as sixth argument. Note that the address of the termination function initially is located in the `rdx` register. Other registers besides `rdx` and `rsp` contain unspecified values. Actually the main point of the `_start` function is to call `__libc_start_main` . So the next action is to prepare for this function.

The signature of the `__libc_start_main` function is located in the csu/libc-start.c source code file. Let's look on it:

```
STATIC int LIBC_START_MAIN (int (*main) (int, char **, char **),
                             int argc,
                            char **argv,
                             __typeof (main) init,
                            void (*fini) (void),
                            void (*rtld_fini) (void),
                            void *stack_end)
```

It takes the address of the `main` function of a program, `argc` and `argv`. `init` and `fini` functions are constructor and destructor of the program. The `rtld_fini` is the termination function which will be called after the program will be exited to terminate and free its dynamic section. The last parameter of the `__libc_start_main` is a pointer to the stack of the program. Before we can call the `__libc_start_main` function, all of these parameters must be prepared and passed to it. Let's return to the sysdeps/x86_64/start.S assembly file and continue to see what happens before the `__libc_start_main` function will be called from there.

We can get all the arguments we need for `__libc_start_main` function from the stack. At the very beginning, when `_start` is called, our stack looks like:

```
+-----------------+
|       NULL      |
+-----------------+
|       ...       |
|       envp      |
|       ...       |
+-----------------+
|       NULL      |
+------------------
|       ...       |
|       argv      |
|       ...       |
+------------------
|       argc      | <- rsp
+-----------------+
```

After we cleared `ebp` register and saved the address of the termination function in the `r9` register, we pop an element from the stack to the `rsi` register, so after this `rsp` will point to the `argv` array and `rsi` will contain count of command line arguemnts passed to the program:

```
+-----------------+
|       NULL      |
+-----------------+
|       ...       |
|       envp      |
|       ...       |
+-----------------+
|       NULL      |
+------------------
|       ...       |
|       argv      |
|       ...       | <- rsp
+-----------------+
```

After this we move the address of the `argv` array to the `rdx` register

```
popq %rsi
mov %RSP_LP, %RDX_LP
```

From this moment we have `argc` cand `argv`. We still need to put pointers to the construtor, destructor in appropriate registers and pass pointer to the stack. At the first following three lines we align stack to `16` bytes boundary as suggested in ABI and push `rax` which contains garbage:

```
and  $~15, %RSP_LP
pushq %rax

pushq %rsp
mov $__libc_csu_fini, %R8_LP
mov $__libc_csu_init, %RCX_LP
mov $main, %RDI_LP
```

After stack aligning we push the address of the stack, move the addresses of contstructor and destructor to the `r8` and `rcx` registers and address of the `main` symbol to the `rdi`. From this moment we can call the `__libc_start_main` function from the csu/libc-start.c.

Before we look at the `__libc_start_main` function, let's add the `/lib64/crt1.o` and try to compile our program again:

```
$ gcc -nostdlib /lib64/crt1.o -lc -ggdb program.c -o program
/lib64/crt1.o: In function `_start':
(.text+0x12): undefined reference to `__libc_csu_fini'
/lib64/crt1.o: In function `_start':
(.text+0x19): undefined reference to `__libc_csu_init'
collect2: error: ld returned 1 exit status
```

Now we see another error that both `__libc_csu_fini` and `__libc_csu_init` functions are not found. We know that the addresses of these two functions are passed to the `__libc_start_main` as parameters and also these functions are constructor and destructor of our programs. But what do `constructor` and `destructor` in terms of `C` program means? We already saw the quote from the ELF specification:

> After the dynamic linker has built the process image and performed the relocations, each shared object gets the opportunity to execute some initialization code. ... Similarly, shared objects may have termination functions, which are executed with the atexit (BA_OS) mechanism after the base process begins its termination sequence.

So the linker creates two special sections besides usual sections like `.text`, `.data` and others:

- `.init`
- `.fini`

We can find them with the `readelf` util:

```
$ readelf -e test | grep init
  [11] .init             PROGBITS         00000000004003c8  000003c8

$ readelf -e test | grep fini
  [15] .fini             PROGBITS         0000000000400504  00000504
```

Both of these sections will be placed at the start and end of the binary image and contain routines which are called constructor and destructor respectively. The main point of these routines is to do some initialization/finalization like initialization of global variables, such as errno, allocation and deallocation of memory for system routines and etc., before the actual code of a program is executed.

You may infer from the names of these functions, they will be called before the `main` function and after the `main` function. Definitions of `.init` and `.fini` sections are located in the `/lib64/crti.o` and if we add this object file:

```
$ gcc -nostdlib /lib64/crt1.o /lib64/crti.o  -lc -ggdb program.c -o program
```

we will not get any errors. But let's try to run our program and see what happens:

```
$ ./program
Segmentation fault (core dumped)
```

Yeah, we got segmentation fault. Let's look inside of the `lib64/crti.o` with `objdump`:

```
$ objdump -D /lib64/crti.o

/lib64/crti.o:     file format elf64-x86-64
```

```
Disassembly of section .init:

0000000000000000 <_init>:
   0:	48 83 ec 08          	sub    $0x8,%rsp
   4:	48 8b 05 00 00 00 00 	mov    0x0(%rip),%rax        # b <_init+0xb>
   b:	48 85 c0             	test   %rax,%rax
   e:	74 05                	je     15 <_init+0x15>
  10:	e8 00 00 00 00       	callq  15 <_init+0x15>

Disassembly of section .fini:

0000000000000000 <_fini>:
   0:	48 83 ec 08          	sub    $0x8,%rsp
```

As I wrote above, the `/lib64/crti.o` object file contains definition of the `.init` and `.fini` section, but also we can see here the stub for function. Let's look at the source code which is placed in the sysdeps/x86_64/crti.S source code file:

```
	.section .init,"ax",@progbits
	.p2align 2
	.globl _init
	.type _init, @function
_init:
	subq $8, %rsp
	movq PREINIT_FUNCTION@GOTPCREL(%rip), %rax
	testq %rax, %rax
	je .Lno_weak_fn
	call *%rax
.Lno_weak_fn:
	call PREINIT_FUNCTION
```

It contains the definition of the `.init` section and assembly code does 16-byte stack alignment and next we move address of the `PREINIT_FUNCTION` and if it is zero we don't call it:

```
00000000004003c8 <_init>:
  4003c8:       48 83 ec 08             sub    $0x8,%rsp
  4003cc:       48 8b 05 25 0c 20 00    mov    0x200c25(%rip),%rax        # 600ff8 <_DYNAMIC+0x1d0>
  4003d3:       48 85 c0                test   %rax,%rax
  4003d6:       74 05                   je     4003dd <_init+0x15>
  4003d8:       e8 43 00 00 00          callq  400420 <__libc_start_main@plt+0x10>
  4003dd:       48 83 c4 08             add    $0x8,%rsp
  4003e1:       c3                      retq
```

where the `PREINIT_FUNCTION` is the `__gmon_start__` which does setup for profiling. You may note that we have no return instruction in the sysdeps/x86_64/crti.S. Actually that's why we got a segmentation fault. Prolog of `_init` and `_fini` is placed in the sysdeps/x86_64/crtn.S assembly file:

```
.section .init,"ax",@progbits
addq $8, %rsp
ret

.section .fini,"ax",@progbits
addq $8, %rsp
ret
```

and if we will add it to the compilation, our program will be successfully compiled and run!

```
$ gcc -nostdlib /lib64/crt1.o /lib64/crti.o /lib64/crtn.o  -lc -ggdb program.c -o program

$ ./program
x + y + z = 6
```

Conclusion

Now let's return to the `_start` function and try to go through a full chain of calls before the `main` of our program will be called.

The `_start` is always placed at the beginning of the `.text` section in our programs by the linked which is used default `ld` script:

```
$ ld --verbose | grep ENTRY
ENTRY(_start)
```

The `_start` function is defined in the sysdeps/x86_64/start.S assembly file and does preparation like getting `argc/argv` from the stack, stack preparation and etc., before the `__libc_start_main` function will be called. The `__libc_start_main` function from the csu/libc-start.c source code file does a registration of the constructor and destructor of application which are will be called before `main` and after it, starts up threading, does some security related actions like setting stack canary if need, calls initialization related routines and in the end it calls `main` function of our application and exits with its result:

```
result = main (argc, argv, __environ MAIN_AUXVEC_PARAM);
exit (result);
```

That's all.

Links

- system call
- gdb
- execve
- ELF
- x86_64
- segment registers
- context switch
- System V ABI

Internal `system` structures of the Linux kernel

This is not usual chapter of `linux-insides`. As you may understand from the title, it mostly describes internal `system` structures of the Linux kernel. Like `Interrupt Descriptor Table`, `Global Descriptor Table` and many many more.

Most of information is taken from official Intel and AMD manuals.

interrupt-descriptor table (IDT)

Three general interrupt & exceptions sources:

- Exceptions - sync;
- Software interrupts - sync;
- External interrupts - async.

Types of Exceptions:

- Faults - are precise exceptions reported on the boundary `before` the instruction causing the exception. The saved `%rip` points to the faulting instruction;
- Traps - are precise exceptions reported on the boundary `following` the instruction causing the exception. The same with `%rip` ;
- Aborts - are imprecise exceptions. Because they are imprecise, aborts typically do not allow reliable program restart.

`Maskable` interrupts trigger the interrupt-handling mechanism only when RFLAGS.IF=1. Otherwise they are held pending for as long as the RFLAGS.IF bit is cleared to 0.

`Nonmaskable` interrupts (NMI) are unaffected by the value of the rFLAGS.IF bit. However, the occurrence of an NMI masks further NMIs until an IRET instruction is executed.

Specific exception and interrupt sources are assigned a fixed vector-identification number (also called an “interrupt vector” or simply “vector”). The interrupt vector is used by the interrupt-handling mechanism to locate the system-software service routine assigned to the exception or interrupt. Up to 256 unique interrupt vectors are available. The first 32 vectors are reserved for predefined exception and interrupt conditions. They are defined in the arch/x86/include/asm/traps.h header file:

```
/* Interrupts/Exceptions */
enum {
    X86_TRAP_DE = 0,    /*  0, Divide-by-zero */
    X86_TRAP_DB,        /*  1, Debug */
    X86_TRAP_NMI,        /*  2, Non-maskable Interrupt */
    X86_TRAP_BP,        /*  3, Breakpoint */
    X86_TRAP_OF,        /*  4, Overflow */
    X86_TRAP_BR,        /*  5, Bound Range Exceeded */
    X86_TRAP_UD,        /*  6, Invalid Opcode */
    X86_TRAP_NM,        /*  7, Device Not Available */
    X86_TRAP_DF,        /*  8, Double Fault */
    X86_TRAP_OLD_MF,    /*  9, Coprocessor Segment Overrun */
    X86_TRAP_TS,        /* 10, Invalid TSS */
    X86_TRAP_NP,        /* 11, Segment Not Present */
    X86_TRAP_SS,        /* 12, Stack Segment Fault */
    X86_TRAP_GP,        /* 13, General Protection Fault */
    X86_TRAP_PF,        /* 14, Page Fault */
    X86_TRAP_SPURIOUS,    /* 15, Spurious Interrupt */
    X86_TRAP_MF,        /* 16, x87 Floating-Point Exception */
    X86_TRAP_AC,        /* 17, Alignment Check */
    X86_TRAP_MC,        /* 18, Machine Check */
    X86_TRAP_XF,        /* 19, SIMD Floating-Point Exception */
    X86_TRAP_IRET = 32,    /* 32, IRET Exception */
};
```

Error Codes

The processor exception-handling mechanism reports error and status information for some exceptions using an error code. The error code is pushed onto the stack by the exception-mechanism during the control transfer into the exception handler. The error code has two formats:

* most error-reporting exceptions format;
* page fault format.

Here is format of selector error code:

```
31                           16 15                                  3   2   1   0
+-------------------------------------------------------------------------------+
|                              |                                    | T | I | E |
|           Reserved           |             Selector Index         | - | D | X |
|                              |                                    | I | T | T |
+-------------------------------------------------------------------------------+
```

Where:

* `EXT` - If this bit is set to 1, the exception source is external to the processor. If cleared to 0, the exception source is internal to the processor;
* `IDT` - If this bit is set to 1, the error-code selector-index field references a gate descriptor located in the `interrupt-descriptor table`. If cleared to 0, the selector-index field references a descriptor in either the `global-descriptor table` or local-descriptor table `LDT`, as indicated by the `TI` bit;
* `TI` - If this bit is set to 1, the error-code selector-index field references a descriptor in the `LDT`. If cleared to 0, the selector-index field references a descriptor in the `GDT`.
* `Selector Index` - The selector-index field specifies the index into either the `GDT`, `LDT`, or `IDT`, as specified by the `IDT` and `TI` bits.

Page-Fault Error Code format is:

```
31                                                              4   3   2   1   0
+-------------------------------------------------------------------------------+
|                                                          |     | R | U | R | - |
|                         Reserved                         | I/D | S | - | - | P |
|                                                          |     | V | S | W | - |
+-------------------------------------------------------------------------------+
```

Where:

* `I/D` - If this bit is set to 1, it indicates that the access that caused the page fault was an instruction fetch;
* `RSV` - If this bit is set to 1, the page fault is a result of the processor reading a 1 from a reserved field within a page-translation-table entry;
* `U/S` - If this bit is cleared to 0, an access in supervisor mode (`CPL=0, 1, or 2`) caused the page fault. If this bit is set to 1, an access in user mode (CPL=3) caused the page fault;
* `R/W` - If this bit is cleared to 0, the access that caused the page fault is a memory read. If this bit is set to 1, the memory access that caused the page fault was a write;
* `P` - If this bit is cleared to 0, the page fault was caused by a not-present page. If this bit is set to 1, the page fault was caused by a page-protection violation.

Interrupt Control Transfers

The IDT may contain any of three kinds of gate descriptors:

* `Task Gate` - contains the segment selector for a TSS for an exception and/or interrupt handler task;
* `Interrupt Gate` - contains segment selector and offset that the processor uses to transfer program execution to a handler procedure in an interrupt handler code segment;
* `Trap Gate` - contains segment selector and offset that the processor uses to transfer program execution to a handler procedure in an exception handler code segment.

General format of gates is:

```
127                                                                             96
+-------------------------------------------------------------------------------+
|                                                                               |
|                                Reserved                                       |
|                                                                               |
+--------------------------------------------------------------------------------
95                                                                              64
+-------------------------------------------------------------------------------+
|                                                                               |
|                               Offset 63..32                                   |
|                                                                               |
+-------------------------------------------------------------------------------+
63                               48 47       46  44   42    39             34    32
+-------------------------------------------------------------------------------+
|                                  |       |  D  |   |     |      |   |   |     |
|       Offset 31..16              |   P   |  P  | 0 |Type |0 0 0 | 0 | 0 | IST |
|                                  |       |  L  |   |     |      |   |   |     |
 -------------------------------------------------------------------------------+
31                                   16 15                                      0
+-------------------------------------------------------------------------------+
|                                      |                                        |
|          Segment Selector            |                 Offset 15..0           |
|                                      |                                        |
+-------------------------------------------------------------------------------+
```

Where

- `Selector` - Segment Selector for destination code segment;
- `Offset` - Offset to handler procedure entry point;
- `DPL` - Descriptor Privilege Level;
- `P` - Segment Present flag;
- `IST` - Interrupt Stack Table;
- `TYPE` - one of: Local descriptor-table (LDT) segment descriptor, Task-state segment (TSS) descriptor, Call-gate descriptor, Interrupt-gate descriptor, Trap-gate descriptor or Task-gate descriptor.

An `IDT` descriptor is represented by the following structure in the Linux kernel (only for `x86_64`):

```C
struct gate_struct64 {
    u16 offset_low;
    u16 segment;
    unsigned ist : 3, zero0 : 5, type : 5, dpl : 2, p : 1;
    u16 offset_middle;
    u32 offset_high;
    u32 zero1;
} __attribute__((packed));
```

which is defined in the arch/x86/include/asm/desc_defs.h header file.

A task gate descriptor does not contain `IST` field and its format differs from interrupt/trap gates:

```C
struct ldttss_desc64 {
    u16 limit0;
    u16 base0;
    unsigned base1 : 8, type : 5, dpl : 2, p : 1;
    unsigned limit1 : 4, zero0 : 3, g : 1, base2 : 8;
    u32 base3;
    u32 zero1;
} __attribute__((packed));
```

Exceptions During a Task Switch

An exception can occur during a task switch while loading a segment selector. Page faults can also occur when accessing a TSS. In these cases, the hardware task-switch mechanism completes loading the new task state from the TSS, and then triggers the appropriate exception mechanism.

In long mode, an exception cannot occur during a task switch, because the hardware task-switch mechanism is disabled.

Nonmaskable interrupt

TODO

API

TODO

Interrupt Stack Table

TODO

Полезные ссылки

Загрузка Linux

- Протокол загрузки Linux/x86
- Параметры ядра Linux

Защищённый режим

- 64-ia-32-architectures-software-developer-vol-3a-part-1-manual.pdf

Управление памятью в ядре Linux

- Заметки @lorenzo-stoakes о подсистеме VM в ядре Linux

Программирование последовательного порта

- Программирование 8250 UART
- Последовательные порты на OSDEV

VGA

- Video Graphics Array (VGA)

Ввод/вывод

- Программирвоание портов ввода/вывода

GCC и GAS

- Типы атрибутов GCC
- Директивы ассемблера

Важные структуры данных

- Определение task_struct

Управление памятью

- Серия статей "NUMA Deep Dive"

Полезные ссылки

- Запуск программ в Linux x86
- Разметка памяти при выполнении программы (32 бита)

Спасибо всем участникам перевода:

@Le0nX