Implementing Threads 2
The problem here is that a terminating thread can’t free its own stack, since it would still be using that stack on return from `free`. 

**Thread Termination**

- **Termination**
  - thread becomes zombie
  - if joinable
    - notify waiter, if present
  - if detached
    - disappear
  - thread can’t do this by itself!
while(zombies) {
    delete(zombie);
}
Thread Yield Details

```c
void thread_yield() {
    if (!queue_empty(runqueue)) {
        enqueue(runqueue, CurrentThread);
        thread_switch();
    }
}
```
Time Slicing

- Periodically
  - current thread forced to do a thread yield

```c
void ClockInterrupt(int sig) {
    thread_yield();
}
```

- Implement ClockInterrupt with VTALRM signal
If a signal is to be dealt with via a signal handler, the kernel arranges so that the handler is invoked much like a subroutine, and that when it returns, the process resumes its normal execution. To accomplish this, the state of the process must be saved before invocation of the handler and resumed on return. A major component of the process state information is the contents of the registers: a typical subroutine does not save and restore all of the registers, since some are allowed to be modified (e.g., the register used to pass back a return value); since a signal handler can be invoked at any point, something must be done to save and restore all registers. We must also deal with masking signals while in the handler: invoking a signal handler causes a specified set of signals (including the one that was just delivered) to be masked off; the process’s original signal mask must be restored on return from the handler.
In this sequence of slides, we show how a signal handler is invoked in Linux. In this slide, our program is executing within its “main line.”
A signal has occurred. This causes the process to be interrupted and the kernel is entered. The process’s registers (including its instruction pointer) are saved on the process’s kernel stack.
The kernel pushes a `struct sigframe` onto the user’s stack, containing a copy of the user’s registers (copied from the kernel stack—it includes the user’s program counter, which points to instruction following the point of interruption), the user’s signal mask, and a return address that points to executable code (also pushed onto the stack). This executable code, which is executed on return from the signal handler, invokes the `sigreturn` system call with the signal mask and the user’s registers as its arguments. The effect of the call is to return to the point of the interrupt with the signal mask restored.
Now the system returns our process to user mode and it resumes execution in the appropriate signal handling routine.
When the signal handler returns, it returns to the executable code that was pushed onto the stack (within \textit{sigframe}) by the kernel. This code, as already mentioned, invokes the \textit{sigreturn} system call, which passes a pointer to the \textit{sigframe} structure back to the kernel.
The `sigreturn` handler in the kernel restores the user’s registers and signal mask from the `sigframe` structure and returns back to the point of interrupt.
End of Digression: Back to Time Slicing

- Periodically
  - current thread forced to do a thread yield

```c
void ClockInterrupt(int sig) {
    // SIGVTALRM is now masked
    thread_yield();
    // thread resumes here
}
```

- Implement ClockInterrupt with VTALRM signal
Setting Up Time Slicing

```c
struct sigaction timesliceact;
timesliceact.sa_handler = ClockInterrupt;
timesliceact.sa_mask = VTALRMmask;
timesliceact.sa_flags = SA_RESTART; // avoid EINTR
struct timeval interval = {0, 1};
    // every .001 milliseconds
struct itimerval timerval;
timerval.it_value = interval;
timerval.it_interval = interval;
sigaction(SIGVTALRM, &timesliceact, 0);
setitimer(ITIMER_VIRTUAL, &timerval, 0);
    // time slicing is started!
```
Masking/Unmasking Signals

```c
sigset_t VTALRMmask;
...
sigemptyset(&VTALRMmask);
sigaddset(&VTALRMmask, SIGVTALRM);
...
sigprocmask(SIG_BLOCK, &VTALRMmask, 0);
...
sigprocmask(SIG_UNBLOCK, &VTALRMmask, 0);
```
void uthread_no_preempt_on() {
    uthread_no_preempt = 1;
}

void uthread_no_preempt_off() {
    uthread_no_preempt = 0;
}

void ClockInterrupt(int sig) {
    if (uthread_no_preempt)
        return;
    ...
}
Limitations of User Threads

- Threads are implemented strictly at user level
  - the OS kernel is unaware of their existence
- What happens if a user thread makes a blocking system call, e.g., `read`?
Multiple Processors

```c
void thread_switch( ) {
    thread_t *NextThread, *OldCurrent;

    NextThread = dequeue(RunQueue);
    OldCurrent = CurrentThread;
    CurrentThread = NextThread;
    swapcontext(&OldCurrent->context, &NextThread->context);
}
```

- How do we employ multiple processors?
  - code merely switches the caller's processor to another thread
- What if the RunQueue is empty?
Solution Sketch

• Introduce idle threads, one for each processor
• Thread calling thread_switch switches to idle thread for its current processor
• Idle thread then switches to first thread on RunQueue, if any
• If RunQueue is empty, idle thread repeatedly checks RunQueue until it’s not empty, then switches to first thread
This code is deceptively simple-looking. Each idle thread starts off by calling `IdleThread_switch()`. It save its context at line 10, and then goes on to repeatedly check the `RunQueue`. When one finds a thread on the `RunQueue`, it switches to that thread’s context in line 14.

Threads calling `thread_switch` first save their registers by calling `getcontext` at line 3. The variable `first` (declared volatile to force gcc not to put it in a register, but to leave it in memory) is set to 1 so that, on the (first) return from `getcontext`, its value is still 1 and the thread continues executing sequentially. At line 7 the call to `setcontext` switches to the registers of the processor’s idle thread. That context was saved in line 10, and thus control continues at line 11, in the context of idle thread, as explained in the preceding paragraph.

When an idle thread switches to a normal thread (by calling `setcontext` at line 14), the normal thread resumes execution starting from where it had last saved its context, at line 3. Thus thread continues at line 4. At this point, `first`, its local variable, now has a value of zero, and thus the thread returns from its original call to `thread_switch`.
MP Mutual Exclusion

- Two sorts
  - spin locks
    - threads wait by repeatedly testing the lock
  - blocking locks
    - threads wait by sleeping, depending on some other thread to wake them up
Note that the CAS instruction is implemented (in hardware) as an atomic instruction: its effect is instantaneous: nothing else can affect memory during the execution of the instruction. In the x86 and x86-64 architectures, the effect of CAS is achieved with the cmpxchng (compare and exchange) instruction with the lock prefix.
void spin_lock(int *spin) {
    while(CAS(spin, 0, 1))
        ;
}

void spin_unlock(int *spin) {
    *spin = 0;
}
Better Spin Lock

```c
void spin_lock(int *spin) {
    while (1) {
        if (*spin == 0) {
            // the mutex was at least momentarily unlocked
            if (!CAS(spin, 0, 1))
                break;  // we have locked the mutex
            // some other thread beat us to it, so try again
        }
    }
}
```
Blocking Locks

```c
void blocking_lock(mutex_t *mut) {
    if (mut->holder != 0) {
        enqueue(mut->wait_queue, CurrentThread);
        thread_switch();
    } else {
        mut->holder = CurrentThread;
    }
}

void blocking_unlock(mutex_t *mut) {
    if (queue_empty(mut->wait_queue)) {
        mut->holder = 0;
    } else {
        mut->holder =
        enqueue(mut->wait_queue);
        enqueue(RunQueue, mut->holder);
    }
}
```

Does it work?

---

Hint: the answer is no.
Note that even this implementation doesn’t work. We must make certain that the spin lock is not unlocked until the blocking thread is no longer running (and, in particular, no longer using its stack).
For details on futexes, avoid the Linux man pages, but look at http://people.redhat.com/drepper/futex.pdf, from which this material was obtained. Note that there’s actually just one futex system call; whether it’s a wait or a wakeup is specified by an argument.

**Futexes**

- Safe, efficient kernel conditional queueing in Linux
- All operations performed atomically
  - `futex_wait(futex_t *futex, int val)`
    - if `futex->val` is equal to `val`, then sleep
    - otherwise return
  - `futex_wake(futex_t *futex)`
    - wake up one thread from futex’s wait queue, if there are any waiting threads
These functions are available on most architectures, particularly on the x86. Note that their effect must be *atomic*: everything happens at once.
void lock(futex_t *futex) {
    int c;
    while (((c = atomic_inc(&futex->val)) != 0)
        futex_wait(futex, c+1);
    }

void unlock(futex_t *futex) {
    futex->val = 0;
    futex_wake(futex);
}
void lock(futex_t *futex) {
    int c;
    if ((c = CAS(&futex->val, 0, 1) != 0)
        do {
            if (c == 2 || (CAS(&futex->val, 1, 2) != 0))
                futex_wait(futex, 2);
            while ((c = CAS(&futex->val, 0, 2)) != 0))
        }
    }

void unlock(futex_t *futex) {
    if (atomic_dec(&futex->val) != 1) {
        futex->val = 0;
        futex_wake(futex);
    }
    }

MP Memory Issues

- Naive view is that all processors in MP system see same memory contents at all times
  - they don’t
This slide illustrates the common view of the architecture of a multi-core processor: a number of processors are all directly connected to the same memory (which they share). If one core (or processor) stores into a storage location and immediately thereafter another core loads from the same storage location, the second core loads exactly what the first core stored.

Unfortunately, as we learned earlier in the course, things are not quite so simple.
Real multi-core processors have L1 caches that sit between each core and the memory bus; there is a single connection between the bus and the memory. When a core issues a store, the store affects the L1 cache. When a core issues a load, the load is dealt with by the L1 cache if possible, and otherwise goes to memory (perhaps via a shared L2 cache). Most architectures have some sort of cache-consistency logic to insure that the shared-memory semantics of the previous page are preserved.

However, again as we learned earlier in the course, even this description is too simplistic.
This slide shows an even more realistic model, pretty much the same as what we saw is actually used in recent Intel processors. Between each core and the L1 cache is a buffer. Stores by a core go into the buffer. Sometime later the effect of the store reaches the L1 cache. In the meantime, the core is issuing further instructions. Loads by the core are handled from the buffer if the data is still there; otherwise they go to the L1 cache, and then perhaps to memory.

In all instances of this model the effect of a store, as seen by other cores, is delayed. In some instances of this model the order of stores made by one core might be perceived differently by other cores. Architectures with the former property are said to have delayed stores; architectures with the latter are said to have reordered stores (an architecture could well have both properties).
In this example, one thread running on one processor is loading from an integer in storage; another thread running on another processor is loading from and then storing into an integer in storage. Can this be done safely without explicit synchronization?

On most architectures, the answer is yes. If the integer in question is aligned on a natural (e.g., eight-byte) boundary, then the hardware (perhaps the cache) insures that loads and stores of the integer are atomic.

However, one cannot assume that this is the case on all architectures. Thus a portable program must use explicit synchronization (e.g., a mutex) in this situation.
Shown on the slide is Peterson’s algorithm for handling mutual exclusion for two threads without explicit synchronization. (The me argument for one thread is 0 and for the other is 1.) This program works given the first two shared-memory models. Does it work with delayed-store architectures?

This example is a solution, employing “busy waiting,” to the producer-consumer problem for one consumer and one producer. It works for the first two shared-memory models, and even for delayed-store architectures. But does it work on reordered-store architectures?

This solution to the producer-consumer problem is from “Proving the Correctness of Multiprocess Programs,” by L. Lamport, IEEE Transactions on Software Engineering, SE-3(2) 1977: 125-143.
The point of the previous several slides is that one cannot rely on expected properties of shared memory to eliminate explicit synchronization. Shared memory can behave in some very unexpected ways. However, it is the responsibility of the implementers of the various synchronization primitives to make certain not only that they behave correctly, but also that they synchronize memory with respect to other threads.
Assume these are run on a two-processor system: why does the two-threaded program on the right run faster than the two-threaded program on the left?
Processors usually employ data caches that are organized as a set of cache lines, typically of 64 bytes in length. Thus data is fetched from and stored to memory in units of the cache-line size. Each processor has its own data cache.
Getting back to our example: we have a two-processor system, and thus two data caches. If \( a \) and \( b \) are in the same cache line, then when either processor accesses \( a \), it also accesses \( b \). Thus if \( a \) is modified on processor 1, memory coherency will cause the entire cache line to be invalidated on processor 2. Thus when processor 2 attempts to access \( b \), it will get a cache miss and be forced to go to memory to update the cache line containing \( b \). From the programmer’s perspective, \( a \) and \( b \) are not shared. But from the cache’s perspective, they are. This phenomenon is known as \textit{false sharing}, and is a source of performance problems.

For further information about false sharing and for tools to deal with it, see http://emeryblogger.com/2011/07/06/precise-detection-and-automatic-mitigation-of-false-sharing-oopsla-11/.