On communication among threads

Runtimes for concurrency and distribution

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Premise – 1

- Concurrency is eminently *collaborative*
  - The threads in a concurrent program hardly are fully independent of one another
    - If they were, the program would be perfectly *parallel*
    - Recall the distinction between concurrency and parallelism!
- Stipulating the communication interfaces allowable among threads is a crucial concern in the design of a concurrent language
  - The chosen model of communication has large impact on the overall quality of the program
    - For efficiency, understandability, maintainability

Premise – 2

- Inter-thread communication can be
  - **Direct**, only involving active entities
  - **Indirect**, mediated by reactive or passive entities
- Classic models
  - **Message passing**, *direct*
    - No sharing: awkward when running on shared memory, but also **very *scalable***
  - **Shared variables**, *indirect*
    - Natural when running on shared memory, but also very risky and **not scalable**
- Before proceeding, make sure you understand how primary memory is organized
  - What can a thread "see", what it cannot

Premise – 3

- Having to synchronize (waiting during execution) to communicate *defeats* parallelism
  - Message passing requires synchronization between sender and receiver
  - Data sharing requires synchronization to serialize data access
  - In either case, waiting may require suspension or spinning
- When data sharing cannot be avoided in a *parallel* system, ***wait-free synchronization*** becomes desirable to contain performance loss
  - Transactional memories can be useful in that case
    - They use concurrency control mechanisms similar to those required for DBs, but realized by HW
    - Consistency (writes are serialized) and isolation (no leaks of partial states) warrant atomicity

Shared variables – 1

- The code fragments (e.g., procedures) that operate on shared variables are termed ***critical sections***
  - Very general definition that makes no assumption on the structuredness of the language
    - Plain code, normal procedures/methods, “special” features
- Concurrent (hence preemptive) access to a critical section may give rise to **data races**
  - Situations where the values assigned to and read from shared variables cannot be predicted
  - A source of *non-determinism*, evil for program verification
    - (We shall see later that, in other cases than critical sections, some degree of non-determinism is desirable)
- The medicine to this risk is **atomicity**

Shared variables – 2

- **Bernstein's condition,** IEEE TREC 15:15, 1966
  - Atomic execution is guaranteed if shared variables that are read and modified by a critical section are *not* modified by any other concurrently executing section of code
- If that condition does not hold, the risk of ***data race*** arises, which may result in *race conditions*
  - Ascertaining the presence of data races in a program is inordinately complex (NP-hard) in the general case
    - R. Netzer and B. Miller, ACM LoPLAS 1:1, 1992

Defeating data races

- The problem has two parts
  - Ensuring that critical sections execute atomically (**P1**)
    - Errors of this type cause **low-level** data races
  - Encapsulating critical sections correctly (**P2**)
    - Errors of this type cause **high-level** data races
- P2-type errors have two ramifications
  - **Non-atomic protection fault**: when a thread's operation on a shared variable is broken up in multiple disjoint partial accesses
  - **Lost-update fault**: when a foreign write to a shared variable occurs between the read and the subsequent functionally-related write of it by one and the same thread

P1-type problem: example – 1

```
// thread A needs to access shared
// variable X
// to this end, it checks whether
// X is free
if (lock == 0) {
   // X is being used
   // try again (busy wait)
}
else {
   // X is free
   // set it to «in use»
   lock = 0;
   <critical section S1(X)>;
   // free X
   lock = 1;
}
```

```
// thread B needs to access shared
// variable X
// to this end, it checks whether
// X is free
if (lock == 0) {
   // X is being used
   // try again (busy wait)
}
else {
   // X is free
   // set it to «in use»
   lock = 0;
   <critical section S2(X)>;
   // free X
   lock = 1;
}
```

Critical sections S1 and S2 are not atomic: why?

P1-type problem: example – 2

```
/* DEPOSIT */

amount = read_amont();
lock(); // this opens
        // a critical section

balance = balance + amount;
interest = interest + rate *
           balance;

unlock(); // this closes
          // a critical sction
```

```
/* WITHDRAW */

amount = read_amount();
if (balance < amount) {
   // notify caller that
   // the operation is denied
}
else {
   balance = balance - amount;
   interest = interest +
              rate * balance;
}
```

`Withdraw` **exposes** `Deposit` to a ***low-level data race***: why?

P2-type problem: example – 1

```
/* Updater Task */

// set status value reading
synchronized (table){
  table[N].value = V;
}

… // do work

// set system status for value N
synchronized (table) {
  table[N].achieved = true;
}
```

In this time span, `table[N]` is not protected

```
/* Monitor Daemon */

synchronized (table){
 if (table[N].achieved &&
   system_state[N] !=
     table[N].value){
    // inconsistent system state
    issueWarning();
   }
}
```

NASA
Remote Agent (1997)
using Java and LISP

A case of ***non-atomic protection fault***: why?

P2-type problem: example – 2

```
/* WITHDRAW */

void withdraw(int amount){
  lock(l);
  int tmp = balance;
  unlock(l);
  if tmp > amount){
    lock(l);
    balance = tmp - amount;
    unlock(l);
  }
}
```

Read access

Write access

```
/* DEPOSIT */

void deposit(int amount){
  lock(l);
  balance = balance + amount;
  unlock(l);
}
```

A case of ***lost-update fault***: why?

Access control fundamentals – 1

- **Exclusion synchronization**
  - When, at any point in time, no more than one thread may have access to a shared resource
    - Access is exclusive
- **Avoidance synchronization**
  - When certain functional preconditions must hold before access can be granted
    - Dependent on the program logic
    - Epitomized by the case of the *bounded buffer*

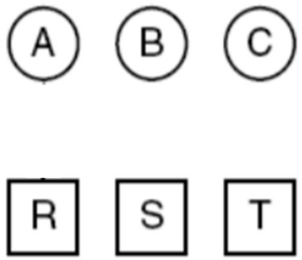
Access control fundamentals – 2

- Synchronization is exposed to risks
  - **Deadlock** or **starvation** (aka lockout)
- Starvation (lockout) occurs when contenders use CPU time *without* making progress
  - As in an unlucky test-and-set situation …
- Deadlock occurs when the involved participants relinquish the CPU and wait indefinitely
  - *Circular-wait* deadlock occurs when 4 conditions hold simultaneously
    1. Mutual exclusion is in use
    2. Resource access cannot be pre-empted
    3. Resource accumulation is allowed with hold-and-wait
    4. The wait condition is circular

Access control fundamentals – 3

- 4 types of reaction to deadlock
  - Ostrich (don't look and hope for the best)
  - Design-time prevention
    - Condition-4 potential can be detected if the participant set is fully and statically known
    - Condition 3 can be defeated forbidding resource accumulation
  - Run-time prevention
    - To combat condition 4, the runtime must stay current of the status of all shared variables (who's holding, who's waiting)
    - Denying access if allowing it risks circular wait
    - Or requiring that access is granted only in a fixed order
  - Run-time detection
    - Oh boy, some threads are not touching the ready queue …

An example of deadlock prevention

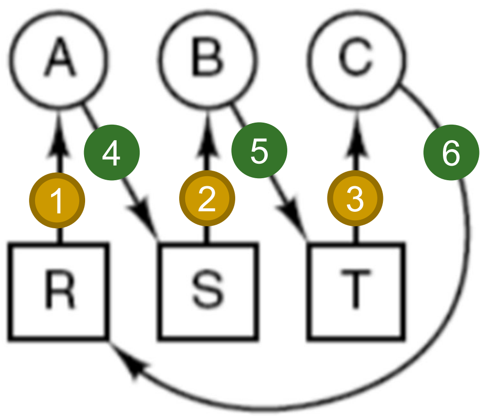


A	B	C
Request R	Request S	Request T
Request S	Request T	Request R
Release R	Release S	Release T
Release S	Release T	Release R

The following interleaving …

1. A requests R
2. B requests S
3. C requests T
4. A requests S
5. B requests T
6. C requests R

… leads to a circular wait



Imagine now that resources could *only* be accessed in a given order (e.g., R, S, T).
In that case, C should request R *before* requesting T …

Access control fundamentals – 4

- Wait time owing to synchronization should be upper bounded
  - Only FIFO queuing ensures that property
    - FIFO policy is (bounded) fair and warrants *liveness*
  - Any other policy, no matter how much common-sense, is exposed to ***starvation***
    - Priority ordering
    - LIFO
    - Urgency

Synchronization solutions – 1

- Good synchronization solutions warrant
  1. Exclusive access
  2. Bounded wait
  3. No assumptions on the behaviour of the execution environment
  4. No threads *outside* of the critical section can influence the access policy

Synchronization solutions – 2

- Regulatory control with a shared variable and strict alternation

```
Thread A ::
while (TRUE) {
  while (turn != 0); // busy wait
  critical_section();
  turn = 1; // alternation
  …
}
```

```
Thread B ::
while (TRUE) {
  while (turn != 1); // busy wait
  critical_section();
  turn = 0; // alternation
  …
}
```

- Defects
  - Busy wait
  - The decision on the alternation is taken *outside* of the critical section
  - Risk of data race on the control variable (not severe)

Synchronization solutions – 3

- **Dekker's algorithm**

```
var flag: array [0..1] of boolean;
turn: 0..1; -- i, j are two threads
repeat
      flag [i] := true;
      while flag [j] do
          if turn = j then
               begin
                    flag [i] := false;
                    while turn /= i  do no-op;
                    flag [i] := true;
               end;
          end if;
      end while;
      critical section
      turn := j;
      flag [i] := false;
      remainder of computation
until false;
```

Busy wait!

Conceived by T.J. Dekker (says E.W. Dijkstra) and later improved (1981).
By setting **flag[i] ← true** , thread **i** requests access. Similarly for thread j.
The value of **turn** arbitrates access between the two threads (i and j).
Can be generalized to more than 2 threads

Synchronization solutions – 4

- **Peterson's algorithm**
  - For pairs of threads
  - Access control logic similar to Dekker's
    - A private flag
    - A shared control variable
  - Exposed to data races if control variable is cached
  - Bounded fair
    - Booking request gives priority to the contender

```
set (flag.mine);
coin := other;
loop
 if (flag.other = clear) continue;
 if (coin = mine) continue;
end loop
// CRITICAL SECTION
clear (flag.mine);
```

Synchronization solutions – 5

```
typedef struct {
    int count;
    queue q; /* queue of threads waiting on this semaphore */
} Semaphore;

void P(Semaphore s)
{
    Disable interrupts;
    if (s->count > 0) {
        s->count -= 1;
        Enable interrupts;
        return;
    }
    Add(s->q, current_thread);
    sleep(); /* re-dispatch */
    Enable interrupts;
}

void V(Semaphore s)
{
    Disable interrupts;
    if (isEmpty(s->q)) {
        s->count += 1;
    } else {
        thread = RemoveFirst(s->q);
        wakeup(thread); /* put thread on the ready queue */
    }
    Enable interrupts;
}
```

Argh!

Who calls these?

The initialization value set to **`count`** determines the type of semaphore:

- `count=1` → binary semaphore
- `count>1` → counting semaphore
- `count=0` → barrier

Leaving the use of **`P(s)`** and **`V(s)`** to the programmer's discipline is risky

The monitor – 1

- An explicit *syntactic structure* (known to the compiler) that encapsulates shared variables and publishes the operations allowed to access them
  - Charles A R Hoare, "*Monitors – An Operating System Structuring Concept*", CACM 17(10):549-557 (1974)
- The shared variable cannot be accessed from outside of the monitor
  - This allows the compiler to assure consistent access control
- It is the calling of monitor operations that triggers access control by the runtime
  - Not the programmer to place locks!

The monitor – 2

- Having a protected shared state allows deciding what to do when that state is not fit for use by a caller that has gained access to it
  - For example, establishing that one cannot write into a shared buffer that is full, and cannot read from a shared buffer that is empty
- The monitor provides ***condition variables*** that can be signalled and waited for
  - Caller is suspended by *waiting* on condition `C` currently false
  - Suspended thread at the top of wait queue is resumed on lock holder *signaling* `C` to have become true

The monitor – 3

```
monitor Container
   condition not-empty := false;
             not-full  := true;
   integer content := 0;

   procedure Insert(prod : integer);
   begin
      if content = N then Wait(not-full);
      <add prod to container>;
      content := content + 1;
      if content = 1 then Signal(not-empty);
   end;

   function Fetch : integer;
   begin
      if content = 0 then Wait(not-empty);
      content :=  content - 1;
      if content = N-1 then Signal(not-full);
      return (<fetch from container>);
   end;

end monitor;
```

```
thread Producer ::
 prod : integer;
begin
   while true do
   begin
      Produce(prod);
      Container.Insert(prod);
   end;
end;
```

```
thread Consumer ::
 prod : integer;
begin
   while true do
   begin
      prod :=
        Container.Fetch;
      Consume(prod);
   end;
end;
```

The monitor – 4

- Calling **`Wait`** on condition variable `Var` blocks the caller when `Var` is false
  - Variable `Var` should describe the resource state
  - The caller (lock holder) relinquishes the CPU and it is placed in a wait queue
  - What happens to the lock at this point?
- Calling **`Signal`** on `Var` releases the thread at the top of the wait queue for `Var`
  - The program's logic decides when `Signal` should be called
  - Which thread gets the lock at this point?
- The compiler makes sure that such calls are atomic and therefore exempt from data races

The monitor – 5

- The monitor concept is vastly better than semaphore-protected critical sections
- But it has defects too
  - The monitor does not let the program decide which the order of calls to it should be at run time
    - The thread that gets there first, access it even if it may have to wait on a false condition variable: big waste!
  - The monitor leaves to the programmer the choice of when to call `Wait` and `Signal`
    - Yes, this is part of the program's logic
    - But the programmer may get it wrong

Java's failed monitor – 1

```
class Monitor{
   private int cont = 0;
   public synchronized void Insert(int prod){
      if (cont == N)
         Block();
      <add prod to container>;
      cont = cont + 1;
      if (cont == 1)
         [this.]notify();
   }
   public synchronized int Fetch(){
      if (cont == 0)
         Block();
      cont = cont - 1;
      if (cont == N-1)
         [this.]notify();
      return(<fetch from container>);
   }
   private void Block(){
      try{[this.]wait();
      } catch(InterruptedException exc) {};}
}
```

```
static final int N = <…>;
static Monitor Container = new Monitor();
// …
Monitor.Insert(prod); // producer
// …
prod = Monitor.Fetch (); // consumer
```

For real?

Java’s failed monitor – 2

- In truth, exclusion synchronization (ES) and avoidance synchronization (AS) are orthogonal concerns
  - ES pertains to access control
  - AS cares that the callers’ operation are consistent with the resource logic
- Java collapses them into a single wait queue
  - What blocked caller does **`notify`**`()` awaken?
    - **`notifyAll`**`()` was invented to do damage control, yielding worse chaos
  - Who gets the lock after **`wait`**`()` and **`notify`**`()`?

Message passing – 1

- Its synchronous variant requires *both* sender *and* receiver to wait for one another
  - In this way, both parties know about the progress state of the other even *without* exchanging data
- As synchronization's wait contrasts parallelism, *asynchronous* message passing becomes attractive
  - Sending is non-blocking
    - The sender delivers to a mailbox and proceeds if there is no receiver yet
    - But then the two parties no longer know about each other's progress
  - Receiving blocks until synchronization ends
    - The receiver that gets there first (no message yet), waits until the sender arrives and delivers

Message passing – 2

- Both variants can be played with to inverse their behaviour
  - Synchronous becomes Asynchronous
    - By placing an intermediary between Sender and Receiver
  - Asynchronous becomes (almost) Synchronous
    - Having Sender await an ack from Receiver
- How do Sender (**S**) and Receiver (**R**) get to know each other?
  - By unique name (of thread, of mailbox)
    - CSP's message passing is synchronous and unidirectional
    - Totally unfit for servers !
  - By type of message / channel at destination

Message passing – 3

- Synchronous communications allow for *bidirectional* data exchange
  - First S to R, then R to S
- Receivers can become servers by exposing multiple bidirectional channels (**entries**)
  - Entries have by-copy `in` and `out` parameters
  - A server exposing multiple entries must specify explicitly which one to service at a given time
- Callers (clients) must name the server and the entry of interest
- Thanks to synchronization, receivers (servers) do **not** need to name their callers
  - This makes the naming relation *asymmetric*

Message passing – 4

- Prefixing specific preconditions (**guards**) to attending to receive calls (entries), allows servers to establish service logic
  - Dijkstra's model of non-deterministic guarded select receive command
    - E.W. Dijkstra, "*Guarded Commands, Nondeterminacy, and Formal Derivation of Programs*", CACM, 18(8):453-457 (1975)

```
select
 Guard_1 => accept Service_1(…);
or
 …
or
 Guard_K => accept Service_K(…);
end select;
```

- Guards are Boolean expressions
  - When they are true (open) the respective receive command (`accept`) is enables on the corresponding channel (entry)
  - When multiple guards are open and calls are pending on the corresponding entry, the choice is non-deterministic