Transaction-Friendly Condition Variables *

Chao Wang, Yujie Liu, and Michael Spear
Lehigh University
{chw412, yul510, spear}@cse.lehigh.edu

Abstract

Recent microprocessors and compilers have added support for transactional memory (TM). While state-of-the-art TM systems allow the replacement of lock-based critical sections with scalable, optimistic transactions, there is not yet an acceptable mechanism for supporting the use of condition variables in transactional programs.

We introduce a new implementation of condition variables, which uses transactions internally, and which can be used from within transactions. Our implementation is compatible with existing C/C++ interfaces for condition synchronization. By moving most of the mechanism for condition synchronization into user-space, our condition variables have low overhead and permit flexible interfaces that can avoid some of the pitfalls of pthread condition variables. Performance evaluation on an unmodified PARSEC benchmark suite shows superior performance for lock-based code. In addition, our transactional condition variables make it possible to replace all locks in PARSEC with transactions.

1. Introduction

The current Draft C++ Transactional Memory Specification [1] makes it relatively easy to replace lock-based critical sections with transactions: one need only replace lock-based regions of code with lexically scoped transactions. One of the most significant obstacles that remains is condition synchronization. Both Vyas et al. [22] and Skyrme and Rodriguez [18] observed this problem in the course of their efforts to transactionalize memcached and luaproc, respectively.

In C++, a condition variable must be associated with a named mutex. A thread must hold the mutex when waiting on the condition variable, and the act of waiting effectively completes one critical section. The thread can then be put to sleep and another thread may safely acquire the lock and modify shared state. If its modifications satisfy the condition, it signals the condition variable. The waiting thread can then be woken, at which point it acquires the mutex and executes the continuation of its critical section.

In existing concurrent programs, the act of breaking atomicity at the point of a thread wait does not compromise correctness: the programmer is responsible for checking and restoring invariants when the thread resumes execution. Thus a straightforward translation of the synchronizing critical section to a pair of transactions is not, itself, a concern. The problem is that a waiting thread must release the lock and put itself to sleep in a single atomic operation; if the sleep is delayed until after the lock is released, then it is possible for an intervening signal operation to "miss" the waiting thread.

Traditionally, this problem is solved by implementing the locking and waiting mechanisms within the operating system. In this manner, the operating system is able to mark the thread as waiting, release the lock, and schedule another thread in a manner that appears atomic with respect to all other threads. A common practice in this case is to relax the guarantees made by the operating system: in Mesa, a signal could accidentally wake more than one thread [4], and in the POSIX specification, a wait operation can return without being paired with a signal. It should be noted that modern code accepts these relaxations, and employs a few simple patterns (i.e., waiting within a while loop) to overcome any potential spurious wake-ups.

In this paper, we present a novel implementation of condition variables that is compatible with both locks and transactional memory. Our work is inspired by prior work by Dudnik and Swift [6], which discusses an extension to the Solaris OS that supports condition variables within a research hardware TM prototype, and AtomCaml [16], which was the first to consider splitting `wait` operations within transactions. The key innovation in our work is that each condition variable is implemented as a transactional queue of per-thread counting semaphores. This design avoids several pitfalls identified by Birrell [3], and makes it possible to implement condition variables portably, without OS modification. The result is *simpler* and *more flexible* than traditional condition variables, does not have a noticeable impact on performance, and is agnostic to the TM implementation (i.e., it is compatible with both hardware and software transactions). Furthermore, our implementation is not prone to certain spurious wake-ups that can occur with condition variables implemented within the OS.

The remainder of this paper is organized as follows. Section 2 presents our condition variable algorithm. Section 3 discusses our implementation in C++. In Section 4, we evaluate our algorithm using both lock-based and transactionalized versions of the PARSEC [2] benchmark suite. In Section 5, we review related work, with a focus on programming models. Section 6 discusses future work and concludes.

2. A Condition Variable Algorithm

Unlike linearizable concurrent data structures, a condition variable does not naively admit a sequential specification: an invocation of the wait method cannot produce a response without it paring with an intervening operation (a signal) by another thread. Nonetheless, we require a specification of the behavior of condition variables before we can describe an algorithm that can be proven correct. To that end, we begin by presenting an abstract specification modeled after that proposed by Birrell et al. [4], and the sequential specification of a lower-level CondVar object. Using this specification, we introduce a generic algorithm and prove that it is both correct and immune to spurious wake-ups. We then discuss synchronization requirements.

2.1 Conventional Specification

In order for a condition variable to be useful, a programmer must be able to reason about the order in which WAIT and NOTIFY opera-

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Algorithm 1: The CondVar Specification

shared states

Q : `Set<Thread>` // waiting threads; initially $\emptyset$

// p refers to the thread that performs the operation

procedure WAITSTEP1()
  $Q \leftarrow Q \cup \{p\}$

function WAITSTEP2() : `Boolean`
  return $p \in Q$

procedure NOTIFYONE()
  if $\exists\, x \in Q$ **then** $Q \leftarrow Q \setminus \{x\}$

procedure NOTIFYALL()
  $Q \leftarrow \emptyset$

Algorithm 2: A Generic CondVar Implementation

shared states

Q : `Set<Thread>` // waiting threads; initially $\emptyset$
$spin_p$: `Boolean` // per-thread flag; initially **false**

procedure WAITSTEP1()
1 $spin_p \leftarrow$ **true**
2 $Q \leftarrow Q \cup \{p\}$

function WAITSTEP2() : `Boolean`
3 **while true do if** $\neg spin_p$ **then return false**

procedure NOTIFYONE()
  // remove from Q an arbitrary element x if exists
4 **if** $\exists\, x \in Q$ **then** $\{Q \leftarrow Q \setminus \{x\}; e \leftarrow$ **true**$\}$ **else** $e \leftarrow$ **false**
  // clear $spin_x$ if some x is removed from Q by last step
5 **if** e **then** $spin_x \leftarrow$ **false**

procedure NOTIFYALL()
  // move all elements from Q to Q'
6 $\langle Q' \leftarrow Q\ ;\ Q \leftarrow \emptyset \rangle$
  // remove some x from Q' and clear $spin_x$
7 **while** $\exists x \in Q'$ **do** $\{Q' \leftarrow Q' \setminus \{x\}; spin_x \leftarrow$ **false**$\}$

tions are performed. Traditionally, this is accomplished by coupling the use of condition variables with mutual exclusion locks.

Birrell et al. [4] proposed a semantics that is widely adopted by many implementations of condition variables. In Birrell's specification, the abstract states of a condition variable object consist of a set Q of waiting threads (initially empty) and a mutual exclusion lock L. The object supports three operations specified as follows:

- A WAIT operation must be invoked within a critical section where L is held by the invoking thread p. The operation consists of two *separate* atomic steps: the first step adds p to Q and releases L atomically; in the second step, the thread is suspended until reaching a state where $p \notin Q$ and L is not acquired, at which point it acquires L and returns.
- A NOTIFY operation atomically removes some non-empty proper subset of threads from Q if Q is not empty.
- A NOTIFYALL operation atomically makes Q empty.

2.2 A Common Specification

Our aim is to produce a common specification of condition variables that is compatible with both locks and transactions. To this end, we need to eliminate the notion of locks from the specification. We also notice that in Birrell's specification, NOTIFY can be simply implemented as a NOTIFYALL, and thus, we remove the former operation from the interface, and replace it with a NOTIFYONE operation to allow removing exactly one thread from the set of waiting threads.

We introduce *atomic sequences* as the foundation of our common specification. An atomic sequence $\langle S \rangle$ is a dynamic sequence of instructions S executed by some thread p, which are enclosed by special beginning and ending instructions. The sequence of instructions S in $\langle S \rangle$ is executed atomically if there is no occurrence of a WAIT operation in S. Atomic sequences are flat-nested, that is, a nested atomic sequence $\langle S_0\ ;\ \langle S_1 \rangle\ ;\ S_2 \rangle$ is semantically equivalent to $\langle S_0\ ;\ S_1\ ;\ S_2 \rangle$.

A WAIT operation can appear only in an atomic sequence. An atomic sequence $\langle S; \text{WAIT}; C \rangle$, where S is the preceding sequence before the first occurrence of WAIT and C is the continuation sequence after the WAIT operation, is semantically equivalent to the following sequence:

$$\langle S\ ;\ Q \leftarrow Q \cup \{p\} \rangle\ ;\ \langle \textbf{assert}\ p \notin Q \rangle\ ;\ \langle C \rangle$$

A NOTIFYONE or NOTIFYALL operation can appear either in an atomic sequence or not. In either cases, a NOTIFYONE is equivalent to $\langle \textbf{if}\ \exists x \in Q\ \textbf{then}\ Q \leftarrow Q \backslash \{x\} \rangle$, and a NOTIFYALL is equivalent to $\langle Q \leftarrow \emptyset \rangle$.

The interface of a CondVar object consists of four operations listed in Algorithm 1: WAITSTEP1, WAITSTEP2, NOTIFYONE, and NOTIFYALL. Intuitively, a WAIT operation (Step1 and Step2) adds the thread to the waiting set and suspends a thread. A NOTIFYONE wakes a thread that has performed WAITSTEP1 on the same CondVar but has not yet been woken, and NOTIFYALL wakes all threads that have performed a WAITSTEP1 on a CondVar but have not yet been woken.

We define the set of *legal* histories by imposing constraints on the set of all sequential histories permitted by the CondVar object in Algorithm 1.

Definition 1. A sequential history of a CondVar object is legal if it satisfies the following:

(1) For every thread, a WAITSTEP1 operation is immediately followed by a WAITSTEP2 in the thread's history.
(2) Every WAITSTEP2 operation returns false.

2.3 A Generic Implementation

Let us now consider an implementation of condition variables that satisfies this specification. We represent each condition variable as a set of thread identifiers, and additionally require per-thread flags. The set stores the identities of all threads waiting on a particular CondVar; the flag is a convenience mechanism that provides a means for decoupling set operations from the instructions that allow a waiting thread to continue.

In this algorithm, a thread performs WAITSTEP1 by setting its flag and then inserting its unique identifier into a particular CondVar's set. To wake a thread, should there be one waiting, a thread uses NOTIFYONE to remove one entry from the set, and then clears the corresponding thread's flag. NOTIFYALL is similar to NOTIFYONE, except it wakes all threads that are sleeping on the CondVar.

We now prove a few properties of this generic implementation. For the proofs, we assume that each line in the code listing is executed as an atomic step. Note that for the while-loops at lines 3 and 7, "executing as an atomic step" means executing one iteration of the loop as an atomic step, including the evaluation of the condition and at most one execution of the loop body. We use the

Algorithm 3: Data Types and Variables

```
QUEUENODE DATA STRUCTURE
  sem      : sem_t        // reference to a semaphore
  next     : QueueNode    // next entry in queue
PER-THREAD VARIABLES
  my_node  : QueueNode    // reference to a queue node
CONDVAR DATA STRUCTURE
  head     : QueueNode    // reference to head of queue
  tail     : QueueNode    // reference to tail of queue
```

notation $p@k$ to denote that thread p is about to execute the step at line k.

The main obligation of the proof is to show that there exists a refinement mapping from the generic implementation to the CondVar specification. The following invariants capture the basic properties of the algorithm, which can be proved together (as one conjunction) by induction over reachable states.

Lemma 2. The following statements hold as invariants:
(1) $p@1 \implies \neg spin_p$
(2) $p@2 \implies spin_p$
(3) $p \in Q \implies p@3 \wedge spin_p$
(4) $p@5 \wedge e \implies x@3 \wedge spin_x$
(5) $p@7 \wedge x \in Q' \implies x@3 \wedge spin_x$

Theorem 3. The generic CondVar implementation is linearizable.

Proof. We define the linearization point of each operation as follows:

- A WAITSTEP1 linearizes at line 2.
- A WAITSTEP1 linearizes at line 3 where it reads $spin_p$ is false.
- A NOTIFYONE linearizes at line 4.
- A NOTIFYALL linearizes at line 6.

□

3. Design and Implementation

We now present a complete implementation of condition variables that is compatible with both locks and transactions. The implementation satisfies the specification from Algorithm 2. We begin by discussing an abstract approach, and then we describe implementation alternatives.

3.1 Data Structures

A practical implementation of condition variables must ensure that threads yield the CPU when they are waiting on the processor, and that there are not arbitrary delays between when a thread is signaled and when it resumes execution. Typically, this is achieved by implementing condition variables as Operating System objects. In contrast we represent each condition variable as a queue in user-space, and implement the per-thread $spin_p$ flags as binary semaphores. The queue stores references to individual threads' semaphores. By initializing the semaphores to 0, we can remove line 1 from Algorithm 2 and implement line 3 as as $sem_p.wait()$. The instances of $spin_x \leftarrow false$ on lines 5 and 7 can each be replaced with as $sem_x.signal()$. Algorithm 3 presents the data structures required to achieve this implementation.

3.2 Algorithm Description

In the interest of generality, we assume a continuation-passing style of execution. The call to WAIT thus takes two parameters: an abstract description of the synchronization context, and the continuation to execute after the thread resumes execution. As we discuss later in this section, our implementation can be adapted to other styles with little effort. Algorithm 4 presents an implementation of WAIT using this interface, and Algorithm 5 presents NOTIFYONE.

Algorithm 4: The Wait algorithm, using continuation passing

```
procedure WAIT(Sync, Cont)
    my_node.next ← nil
    // Insert thread's semaphore into CondVar's queue
    BEGINTRANSACTION ()
    if tail = nil and head = nil then
        head ← tail ← my_node
    else
        tail.next ← my_node
        tail ← my_node
    ENDTRANSACTION ()
    // Break atomicity by completing enclosing synchronized block
    Sync.End()
    // Wait for a signal
    my_node.sem.wait()
    // Invariant: my_node no longer in queue
    // Execute continuation using same synchronization mechanism
    Sync.Begin()
    Cont.execute()
    Sync.End()
```

Algorithm 5: The NotifyOne algorithm

```
procedure NOTIFYONE()
    BEGINTRANSACTION ()
    // If queue not empty, dequeue head element
    sn ← head
    if sn = nil then
        return
    if head = tail then
        head ← tail ← nil
    else
        head ← head.next
    // Wake the thread when we exit from outer transactional scope
    REGISTERHANDLER (sn.sem.signal())
    ENDTRANSACTION ()
```

We expect the WAIT algorithm to be called from an active synchronization context. That is, $Sync$ should refer to a mutual exclusion lock that is held by the caller, or a transaction that is being executed by the caller. (We defer discussion of nested critical sections until Section 3.4). The thread uses a transaction to enqueue its unique node into the CondVar's queue. The use of transactions provides generality and safety: since both WAIT and NOTIFYONE use transactions to access the queue, both methods can be called from any combination of lock-based code, transactional code, and even unsynchronized code without risking data races on the queue. Strictly speaking, if the CondVar methods are always called from the same synchronization context (locks or transactions), this inner transaction is not necessary.

Once the thread has enqueued its semaphore, it then completes its caller's synchronization block, by either releasing locks or committing the transaction. At this point, we know that descheduling of the caller cannot lead to deadlock: it does not hold resources that are required by another thread. Thus it is safe for the thread to wait on its semaphore. Once the semaphore is signaled, the thread will awake, and execute the continuation ($Cont$) in a synchronized manner, in keeping with the synchronization description present in $Sync$. In comparison with Algorithm 2, we see that the only changes are (a) introducing a synchronization context, and (b) replacing spin-waiting on per-thread flags with the use of per-thread semaphores. Note, too, that by explicitly ending one synchronization context and then instantiating another, we can be sure that there is no active hardware or software transaction at the time of the call to $sem.wait()$. Without this guarantee, hardware transactions would abort, due to the system call.

The behavior of NOTIFYONE is simple: using a transaction, the caller removes exactly one element from a nonempty queue, and schedules a signal operation on that element's semaphore. As with WAIT, the use of a transaction ensures race freedom even in the case of naked notifies (i.e., when NOTIFYONE is called from an unsynchronized context). One subtletly is that we use an "onCommit" handler to schedule the semaphore signal to occur when the transaction commits. When NOTIFYONE is called while a lock is held, or from an unsynchronized context, the signal will happen immediately after line 9 completes. However, if NOTIFYONE is called from a transaction, then a waiting thread will not be woken until the caller's outermost transaction commits. From the perspective of Mesa-style semantics, there is no harm in this approach; the wake-up operation can delay. However, by delaying the operation, we can be sure that (a) there is no wake-up caused by a transaction that ultimately does not commit, and (b) there is no attempt to call $sem.signal()$ from an active hardware transactional context. As with WAIT, such a call would cause the hardware transaction to abort and restart in software mode. Note, too, that the current GCC TM implementation maintains the necessary data structures to allow a hardware transaction to store onCommit handlers and run them after transaction commit.

3.3 Supporting NotifyAll

Adding NOTIFYALL support is relatively straightforward. We need only dequeue all elements from the CondVar's queue, and then schedule each element's semaphore to be signaled. An implementation appears in 6.

Algorithm 6: The NotifyAll algorithm

```
procedure NotifyOne()
    BeginTransaction ()
    // If queue not empty, dequeue all elements
    sn ← head
    if sn = nil then
        return
    head ← tail ← nil
    // Wake all threads when we exit from outer transactional scope
    while sn ≠ nil do
        RegisterHandler (sn.sem.signal())
        sn ← sn.next
    EndTransaction ()
```

The principal burden of this algorithm is to ensure that accesses to a queue node's $next$ pointer do not race with nontransactional accesses on line 1 of WAIT. Note that there is a form of privatization taking place: once a thread's node is removed from the queue, there should be no references to the node from any thread other than the node's owner; otherwise, the unsynchronized write on line 1 of WAIT would not be correct.

In the case of NOTIFYALL, we have the guarantee that all accesses to $next$ fields are performed within a transaction. In order for these elements to be accessible to the thread, the element owner must have committed a transaction on line 8 of WAIT, and there cannot have been an intervening NOTIFYONE or NOTIFYALL that removed that thread's node from the queue. Thus it is impossible for the waiting thread to have reached line 11 of WAIT, and no race is possible.

3.4 Specializing Wait for Different Synchronization Contexts

It should first be noted that the availability of a continuation provides a nice optimization: since the continuation represents the code that must execute between when WAIT returns and when the calling context's synchronized region appears to complete, we can elide lines 11–13 in the case where the continuation is empty (i.e., where WAIT is the last instruction in a critical section). When called from a lock-based critical section, this optimization prevents a lock acquire and release pair, decreasing latency and reducing contention on the lock.

On the other hand, when called from a lock-based critical section, the use of a continuation can be avoided. Suppose there was no continuation parameter, and that lines 12 and 13 were elided. In this setting, the behavior of our CondVar, when called from a lock-based critical section, would be indistinguishable from pthread or C++11 condition variables: the caller would execute the continuation upon returning from the call to WAIT, and would do so using the same synchronization mechanism (the same lock) as was in use at the time of the call to WAIT.

This situation can be generalized to a nested monitor environment, in which several locks are held. If $Sync$ stores references to all locks held at the time of the call to WAIT, then all locks can be released (in any order) on line 9, and then can be re-acquired (presumably in order from outermost to innermost [24]) on line 11. As with the single-lock setting, the use of a continuation is not required.

When called from a transactional context, our implementation may require changes to the TM run-time system if a continuation style is not used. In pure-hardware transactions, where the hardware is responsible for thread checkpointing [10, 11], transactions need not be lexically scoped, and thus it is possible to implement our mechanism, without continuations, by simply invoking the hardware primitives to end and begin transactions. Note that in the case of flat nesting, WAIT must record the nesting depth on line 9, and then create an appropriate transaction nest on line 11. Failure to do so will result in the innermost transaction of the loop nest committing its state, and the remaining transactional scopes of the continuation executing nontransactionally.

For software transactions, it is relatively straightforward to allow a transaction to commit early (line 9). However, it is not as easy to open a new transactional context on line 11, return to the caller, and then execute the continuation. For relaxed software transactions, and for atomic software transactions that do not call `cancel` in the continuation, it is possible to run the continuation irrevocably [20, 23], in which case checkpointing is not necessary, and rollback is not possible. As with hardware transactions, the nesting depth will need to be preserved. However, for long-running continuations, such an approach could greatly impede scalability. An alternative is to perform aggressive stack checkpointing as part of line 11, so that the underlying TM can support transactions that are not lexically scoped. We do not advocate this as an interface available to the programmer, only as an option available to the run-time system for the sake of supporting efficient condition synchronization.

For completeness, we note that another option is to require the continuation to be empty. In this case, WAIT lines 9 and 11-13 are elided, and line 10 is deferred via REGISTERHANDLER. For transactions in which the WAIT operation is in a shallow scope relative to the transaction boundaries, the code rewriting that is required for non-empty continuations should be minimal. Furthermore, the runtime can easily provide dynamic checking: after line 10, a flag can be set, which is cleared on transaction commit. Transactional loads and stores would then check the flag before accessing shared memory, and throw an exception if the flag is set. In this manner, the runtime can prohibit memory accesses between the WAIT and the commit of the (rewritten) enclosing transaction.

3.5 Implementation Properties

Our implementation provides the following properties and guarantees to programmers:

Yielding The history of monitors extends back to a time when uniprocessors were prevalent. Even with multicore CPUs, multiprogramming and oversubscription of threads necessitate support for descheduling a waiting thread, and running another thread on the same CPU. Any practical implementation of condition variables must ensure that upon reaching line 3, the calling thread is put to sleep, and also that the delay between when `notify` is called, and when the corresponding thread awakes, is minimal. Clearly, our generic algorithm fails in this regard, as it uses a busy wait loop. Even replacing the busy wait with a call to $sched\_yield$ would not suffice, as it would not guarantee quick wake-up after a notification. However, our use of semaphores addresses this requirement.

Deterministic Wake-Up Semantics In Hoare's work [9], the set associated with a condition variable is explicitly stated to be a queue. Furthermore, a NOTIFYONE operation (there was no NOTIFYALL) was required to be performed while holding a lock, and immediately transferred the lock to the thread at the head of the queue. Mesa, on the other hand, delayed signals until the notifier reached the end of the critical section. This delay, and the absence of an explicit hand-off of the lock, allowed for higher performance at the cost of weaker semantics: when thread a signaled thread b, there was no mechanism to prevent some other thread c from entering the monitor after a completed but before b resumed. This property is shared by our implementation: When a NOTIFYONE pairs with a WAIT operation, there is no bound on the delay between lines 10 and 11.

On the other hand, the use of a generic set, rather than a queue, matches the C++11 and pthread specifications. Thus it is possible that NOTIFYONE may wake any waiting thread, without regard for which thread began waiting first. Scherer and Scott argued that both stack (LIFO) and queue (FIFO) semantics are sometimes advantageous [17], particularly with respect to caching. Our relaxed (i.e., Mesa) semantics for condition variables, coupled with the user-space implementation described in Section 3, allow for arbitrary thread selection policies, with FIFO as the default. Indeed, since the set is in user-space, it is possible to provide a NOTIFYBEST operation, which traverses the set and selects the best thread to wake (possibly using priority, or an additional parameter provided to the WAIT operation to describe the predicate upon which each thread is waiting).

Spurious Wake-Ups Our specification does not allow spurious wake-ups. That is, a call to WAIT cannot return unless it matches with exactly one notifying operation. In our design, the act of putting a thread to sleep need not be atomic with the linearization of its upper half. That being the case, we do not require custom OS support.

In contrast, both the C++11 and pthread specifications allow for a call to WAIT to return even in the absence of a subsequent NOTIFYONE or NOTIFYALL. This relaxation of the specification appears to be a consequence of how operating systems implement condition variables, and in particular how they respond to interrupts that arrive while a call to `wait` is in the midst of transitioning between user-space and kernel execution. In these systems, it must be assumed that *any* call to WAIT may simply return, without a matching signal. Thus even in programs whose logic prevents oblivious wake-ups, WAIT should be called from a loop in order to detect spurious wake-ups. In contrast, our specification does not allow such spurious returns from the WAIT method.

Oblivious Wake-Ups The addition of NOTIFYALL to monitors arose from a common usage pattern in which several threads wait on different predicates. Since traditional condition variables maintain the set of waiting threads in the operating system, it is not possible for NOTIFYONE to know which thread to wake (or even if any thread is sleeping). Consequently, a NOTIFYONE might wake the "wrong" thread, which then must call NOTIFYONE before putting itself back to sleep. The solution, NOTIFYALL, wakes all threads sleeping on a condition variable. When more than one thread is sleeping, we refer to these as "oblivious" wake-ups, since they wake up all of a CondVar's waiting threads, regardless of wether the predicates upon which they depend have been satisfied.

Since the possibility of spurious wake-ups already necessitates that all threads double-check program data upon return from WAIT, allowing oblivious wakeups imposes no burden on the programmer in the general case. Given the absence of spurious wake-ups, our implementation only requires double-checking after a WAIT for specific (general) patterns. In particular, single-producer/single-consumer patterns, which do not cause oblivious wake-ups, are simpler to implement given our stronger specification.

4. Evaluation

In this section, we evaluate the performance of our implementation of transaction-safe condition variables. We seek to answer two main questions, one quantitative and the other qualitative:

- What is the overhead of these condition variables, versus pthread condition variables, in lock-based code?
- What anomalies arise when using these condition variables from transactions?

4.1 Experimental Platforms

We performed experiments using two machines. "Westmere" experiments were performed on a 6-core/12-thread Intel Xeon X5650 CPU running at 2.67GHz; "Haswell" experiments used a 4-core/8-thread Intel Core i7-4770 CPU running at 3.40GHz. Both machines were running Ubuntu 13.04 with kernel version 3.8.0. All benchmarks were compiled using an experimental version of GCC, version 4.9.0. The compiler was configured to use its ml_wt software TM algorithm on Westmere, and to use its HTM algorithm on Haswell. All code was compiled at -O3 optimization level, and experiments are the average of five trials. Variance was uniformly low.

4.2 Benchmarks

We evaluated the performance of our condition variable library using eight benchmarks: facesim, ferret, fluidanimate, streamcluster, bodytrack, x264, raytrace and dedup. These benchmarks are from the PARSEC benchmark suite [2]. Of the 16 benchmarks in PARSEC, three are "network" versions of other benchmarks within PARSEC, and five benchmarks (blackscholes, freqmine, swaptions, vips and canneal) do not use condition variables. We evaluate the remaining eight benchmarks. These benchmarks are representative of the general conditional synchronization patterns that are used widely in current shared-memory multi-threaded programs. We describe these benchmarks and their parallelization and condition synchronization patterns as below.

- **facesim** computes the animation of an input modeled face by simulating its underlying physics. It uses condition variables to implement implements a dynamic and load-balanced task queue that can be employed by a group of working threads. The main program adds tasks to each task queue and waits for the completion of these tasks by the working threads.

- **ferret** is a benchmark for content-based similarity search. To process input data (i.e., images), ferret uses a pipeline that contains 6 stages, each stage containing a thread pool and a job queue. From the perspective of condition synchronization, this benchmark represents a pipelined multi-producer, multi-consumer problem.
- **fluidanimate** simulates incompressible fluid for interactive animation. Condition synchronization is only used to implement a barrier, in place of `pthread_barrier`.
- **streamcluster** solves an online clustering problem. Similar to fluidanimate, streamcluster uses condition variables to implement a barrier. It also employs condition variables to allow a master thread to distribute work in a master/slaves pattern.
- **bodytrack** is a computer vision application that can track the 3-D pose of a human body through a series of images. Condition variables are used to implement three synchronization facilities: a barrier, a multi-threaded synchronization queue, and a persistent thread pool.
- **x264** is an H.264/AVC video encoder. Like `ferret`, `x264` also represents a pipeline model. However, `x264` has as many pipeline stages as input frames. Each thread encodes one frame at a time and all threads work in parallel. The use of condition variables in `x264` is to coordinate threads in the encoding process and the threads waiting for reference frames.
- **raytrace** is a renderer that generates animated 3D scenes. Multiple threads use a multi-threaded task queue, which employs condition variables.
- **dedup** compresses data streams via a 5-stage pipeline, where each stage employs a queue. Condition variables are used in two settings: the per-stage queues, and a coordination mechanism between worker threads and the (serial) output thread.

Of these benchmarks, `fluidanimate`, `streamcluster` and `bodytrack` use condition variables in place of pthread barriers. Strictly speaking, these uses are not necessary. We measure the condition variable-based barrier nonetheless. All benchmarks except `facesim` and `fluidanimate` can run with any number of threads. `facesim` can only run with a number of threads designated by its input file, while `fluidanimate` can only run with a power-of-2 number of threads. All benchmarks were tested with their largest available inputs: for `facesim`, we use input "sim_dev", for others, we use input "native".

4.3 Software Systems Compared

We compare four alternatives. First, we use the baseline PARSEC benchmarks as a **baseline**. This implementation uses pthread locks to protect critical sections, and pthread condition variables for condition synchronization. Second, we evaluate our implementation in an **ideal** setting. In the ideal setting, all calls to WAIT, NOTIFYONE, and NOTIFYALL are made from within lock-based critical sections. In this case, our implementation does not require the use of transactions to protect the internal queue; it is not a realistic implementation, but allows us to assess overheads. Third, we implemented a **general** variant of condition variables. The general implementation uses transactions internally to protect the condition variables' internal queues, and thus is compatible with code that calls CondVar methods from lock-based, transactional, or unsynchronized contexts. By comparing the general, ideal, and baseline algorithms, we can determine whether (a) our mechanism offers competitive performance for lock-based code relative to the current state-of-the-art, and (b) whether the use of transactions in the implementation creates unacceptable overheads. Note that on the "Westmere" machine, the internal implementation uses software transactions, but on "Haswell" the internal implementation employs hardware TM.

Benchmark	Total Txns	Txns that Wait	Txns Refactored
facesim	9	2	0
ferret	3	2	2
fluidanimate	9	2(2*)	2(2*)
streamcluster	7	3(2*)	2(2*)
bodytrack	9	2(1*)	2(1*)
x264	4	1	0
raytrace	14	4(1*)	0
dedup	10	3	3
TOTAL	65	19(6*)	11(5*)

Table 1: Synchronization characteristics of PARSEC source code. Numbers in parenthesis indicate calls to cond_wait used in the barrier implementation.

Our fourth comparison point, **tm**, replaces all locks in the eight benchmarks with transactions, and uses our general implementation of condition variables. To ensure a consistent interface, we opted not to use the continuation passing style, and instead to use manual refactoring to split transactions at the point of a WAIT (see Section 3.4). Table 1 shows that this effort was minimal.

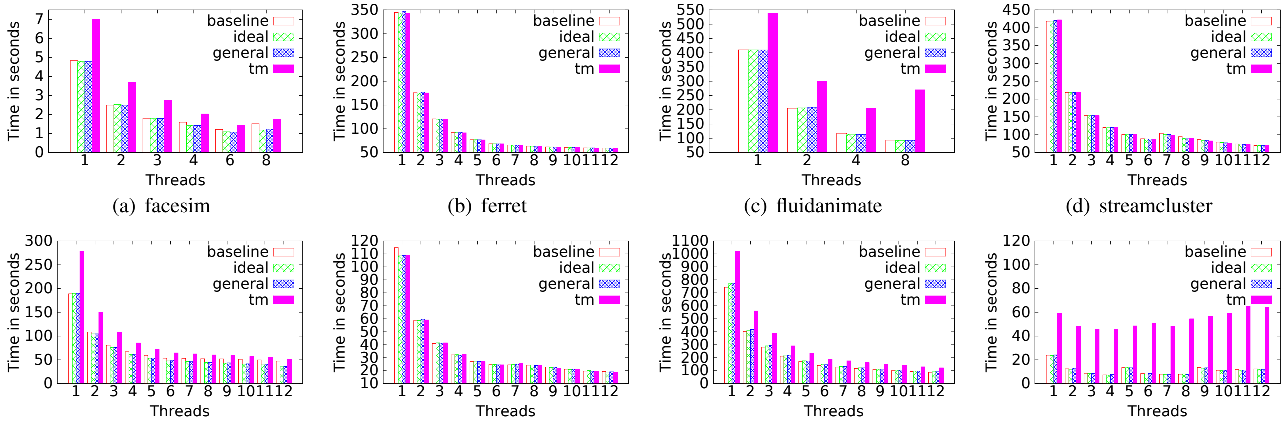
4.4 Performance

Figures 1 and 2 show the performance of PARSEC benchmarks on the Westmere and Haswell machines. Figure 3 presents the geometric mean speedup of each software system, versus the baseline (pthread) implementation of condition variables. We discuss key results below.

Cost of Semaphores On both machines, the ideal and baseline algorithms have very close performance. On Westmere, the ideal algorithm delivers noticable speedups on facesim and bodytrack, and otherwise performs identically to the pthread implementation of condition variables. This indicates that our use of semaphores alone does not impose a poerformance penalty versus an operating system implementation of condition variables. We suspect that the improvements on facesim and bodytrack are a consequence of fewer system calls, since NOTIFYONE need not enter the operating system if the user-space queue is empty. On Haswell, the effect is more muted. Bodytrack again shows an improvement, particularly at high thread counts, but this is offset by additional overhead on raytrace. Taken as a whole, these measurements show that simply using semaphores instead of OS condition variables does not create a performance penalty.

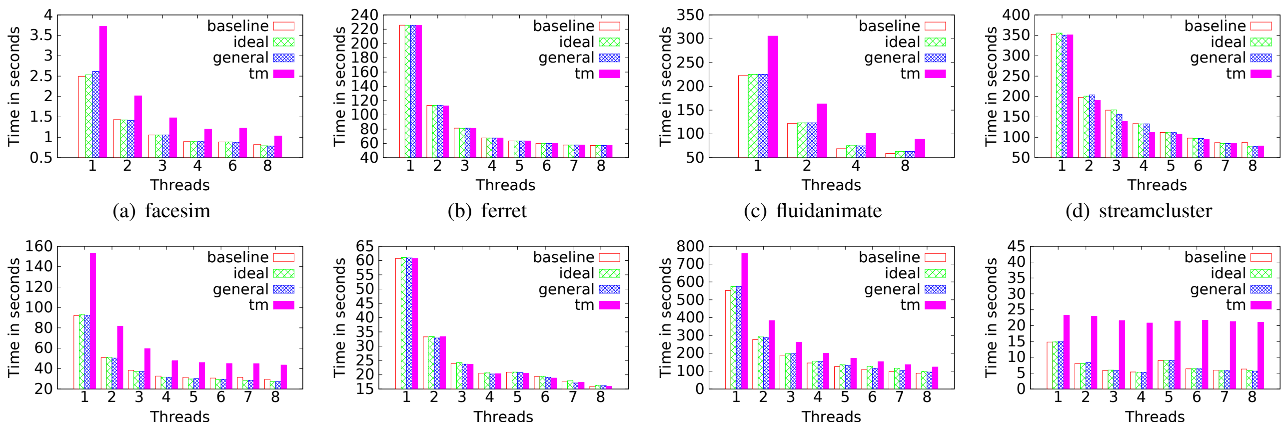
Cost of Condition Variable Implementation The ideal algorithm is not safe for general-purpose code, however, since it assumes that it will be called from a synchronized context. By comparing the ideal and general bars, we can assess the cost of synchronization. Note that we exclusively use TM to synchronize accesses to condition variables' internal data structures. Thus on Westmere, the difference between ideal and general can be attributed to software TM overheads, and on Haswell, the difference is due to hardware TM overheads. The transactions themselves are small, and should not produce conflicts. On both platforms, the overhead is negligible.

Transactionalized PARSEC The tm bars represent the first time that these PARSEC benchmarks have been transactionalized on a real-world system, since no prior work has provided support for condition variables. The benchmark performance roughly falls into three categories. First, on streamcluster, ferret, and x264, performance is roughly equivalent to baseline when all locks are replaced with transactions. Second, in facesim, fluidanimate, bodytrack, and raytrace, we see that performance shows the same tendencies as for lock-based code, but with higher overhead. This should not be a



(a) facesim (b) ferret (c) fluidanimate (d) streamcluster
(e) bodytrack (f) x264 (g) raytrace (h) dedup

Figure 1: Westmere performance



(a) facesim (b) ferret (c) fluidanimate (d) streamcluster
(e) bodytrack (f) x264 (g) raytrace (h) dedup

Figure 2: Haswell performance

surprise: naive transactionalization should not be expected to outperform carefully-tuned lock-based code. We leave as future work investigation into the sources of these overheads.

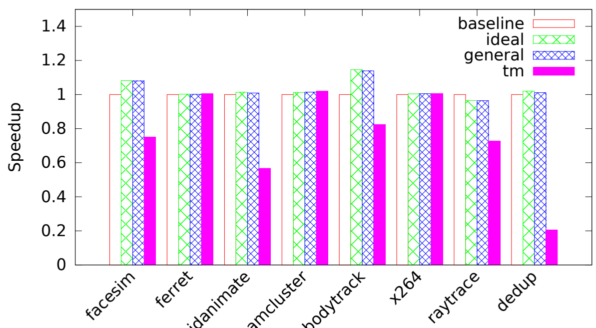
The final benchmark, dedup, exhibits virtually no scaling. In dedup, there is a critical section that performs I/O within a relaxed transaction. These relaxed transactions cannot run in parallel with any other transactions, and thus during I/O, there is no concurrency. While this situation has long been expected by the research community, dedup provides a concrete data point.

Summary Figure 3 summarizes these results: the ideal and the general algorithm only impose negligible performance degradation across all benchmarks, on both machines. The best case speedup, on `bodytrack`, reaches 131.8% on the Westmere, and 110% on Haswell. While there is clearly more work to be done before TM performs as well as locks for PARSEC, the road ahead should be much clearer now that it is possible to transactionalize these 8 benchmarks. Regarding the tm bars, we encourage the reader to treat the results as qualitative: at long last, condition variables can be supported when transactionalizing legacy code, and for many condition synchronization patterns, such integration is seamless and does not impair performance.

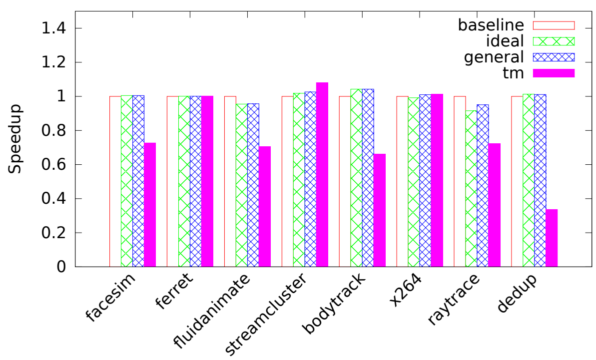
5. Related Work

Research into condition synchronization mechanisms for transactions covers a wide spectrum. On one side, there are efforts, like ours, to improve the implementation of condition variables and/or provide transaction-safe condition variables. On the other side of the spectrum are efforts to craft new alternative condition synchronization mechanisms, which bear little resemblance to traditional monitors and condition variables. While our work is the first to explore condition synchronization in a manner that is compatible with both commodity hardware TM and the C++ TM specification, its relationship with prior work is complex. We highlight salient foundational and related work below.

Our work to replace condition variables with semaphores bears a similarity to efforts undertaken by Birrell [3]. That work attempted to create condition variables for a general-purpose operating system (Win32) using only semaphores. A key consideration was whether it would be possible to implement each condition variable with a constant number of semaphores. Many corner cases arose, which ultimately led to the creation of first-class condition variables in later versions of Win32 operating systems. Birrell's work preceded widespread language-level support for thread-local variables, and thus did not consider the alternative we pro-



(a) Westmere



(b) Haswell

Figure 3: Geometric mean speedup versus baseline

pose, of using per-thread semaphores instead of per-condition variable semaphores.

In another closely related project, Dudnik and Swift [6] explored the hardware and operating system mechanisms required to make transactions compatible with the Solaris operating system's implementation of condition variables. Their work considered a robust hardware TM that had internal support for onCommit handlers, and emphasized compatibility with a legacy implementation of condition variables. As a result, the main considerations dealt with supporting system calls from within a hardware transaction, and making operating system mechanisms compatible with transactional synchronization. In contrast, our work moves all but the most bare-bones scheduler interaction into user-space, and does so in a manner that is agnostic to the TM implementation and operating system. Our work thus generalizes and simplifies this earliest effort at transaction-safe condition variables.

Smaragdakis et al. proposed punctuated transactions as a means of handling I/O and condition synchronization [19]. A punctuating action commits one transaction; user code can then run non-transactionally, after which the continuation executes in a transaction. The continuation must contain programmer-supplied code to verify or restore any invariants that were violated during the period between the two "halves" of the transaction. This model has largely been ignored by the TM community for its complexity; however, the notion of restoring invariants upon resumption of a continuation is precisely the model used for monitor-based synchronization. Our implementation of WAIT can thus be thought of as a specialization of punctuated transactions, in which the only code between the punctuated halves is a $sem.wait()$. It should be straightforward to generalize our implementation to support other forms of punctuation, in a manner that is compatible with both hardware and software transactions.

Efforts to synchronize transactions without splitting them into separate atomic blocks can be separated into two categories: conflict-based coordination, and group commit. In the former category, Harris and Fraser [7], and later the X10 group [5], suggested a Conditional Critical Regions style of synchronization. In this model, the read-only prefix of a transaction determines if a predicate holds, and if not, the transaction aborts and retries. When the predicate holds, the continuation runs in the same context as the predicate test, as a single atomic transaction. To optimize this model, a transaction may make visible the locations it read to compute the predicate, so that it can yield the CPU. The transaction is woken by another thread, after that thread's transaction changes any of the locations upon which the sleeping thread's predicate depends. Harris et al. later extended this approach to a "retry" construct [8], in which transactions can, at any time, determine that a predicate does not hold. At such a point, the transaction may "retry", which rolls back the transaction's effects, makes the transaction's read set visible, and then yields the CPU until some other transaction commits a write to a location that the sleeping transaction had tried to read. Spear et al. [21] later showed that it is possible to manage retry-based read and write set tracking in a manner that is orthogonal to the underlying TM implementation. However, no existing hardware TM systems that can support this mechanism: software instrumentation is currently required.

Lastly, Luchangco and Marathe [14] and Lesani and Palsberg [12] proposed mechanisms for synchronizing transactions via group commit. In these models, which can be extended to more closely resemble condition variables [15], an object mediates dependencies between transactions, and certain interactions (such as one transaction waiting and another signaling) create a requirement for the transactions to commit or abort atomically with each other. As with retry, this mechanism is not compatible with hardware TM: since current hardware TM proposals implement one-phase commit, they cannot ensure that two transactions commit or abort together [13]. As a result, coordinating transactions must execute in software on modern commodity hardware TM.

6. Conclusions and Future Work

This work introduces an implementation of the legacy interface for condition variables that is compatible with transactions. We make it possible, for the first time, to replace locks with transactions in existing software, even when those locks are used for both mutual exclusion and condition synchronization. In experiments on the PARSEC benchmark suite, we showed that the overhead of our mechanism relative to pthread condition variables is negligible, and that the ability to make condition synchronization compatible with transactions allows the discovery of performance anomalies when transactionalizing highly-tuned lock-based code.

Despite this improvement to the state-of-the-art, there is much work that remains, particularly with regard to programming models. In our work, we focus on lock-based code, and thus we do not need to concern ourselves with the use of the "transaction_cancel" construct. Suppose, however, that following the return of a WAIT, the continuation attempts to cancel itself. In this case, it is not clear what should happen: should the outer scope be canceled, in which case a NOTIFYONE might be lost? Should only the continuation be canceled? Real-world uses of both cancellation and transactional condition synchronization are needed before a preferred approach can be known. Indeed, the best approach might be to use a mechanism like $retry$ instead. In that case, a significant challenge will be to allow uninstrumented hardware transactions to run concurrently with a retrying transaction, and to be able to call retry themselves.

Despite these challenges, we are confident that our work will enable more widespread use of transactions. In particular, our effort to make condition variables compatible with hardware TM and the C++ specification ensures that programmers can transactionalize more legacy code, and write new transactional code that uses familiar programming idioms.

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