Atomic RMI 2: Highly Parallel Pessimistic Distributed Transactional Memory

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Abstract

Distributed Transactional Memory (DTM) is an emerging approach to distributed synchronization based on the application of the transaction abstraction to distributed computation. DTM comes in several system models, but the control flow model (CF) is particularly powerful, since it allows transactions to delegate computation to remote nodes as well as access shared data. However, there are no existing CF DTM systems that perform on par with state-of-the-art systems operating in other models. Hence, we introduce a CF DTM synchronization algorithm, OptSVA-CF. It supports fine-grained pessimistic concurrency control, so it avoids aborts, and thus avoids problems with irrevocable operations. Furthermore, it uses early release and asynchrony to parallelize concurrent transactions to a high degree, while retaining strong safety properties. We implement it as Atomic RMI 2, in effect producing a CF DTM system that, as our evaluation shows, can outperform a state-of-the-art non-CF DTM such as HyFlow2.

Index terms—Transactional memory, concurrency control, pessimistic TM, irrevocable transactions, control flow

1 Introduction

Wrangling with concurrency in distributed computing is difficult. Yet, in the era of cloud computing, when anything from simple text editing to big data storage are delegated to remote services, this task is nigh unavoidable. Since application programmers have enough to worry about without delving into the details of distributed computing and networking, these details are comprehensively abstracted away and hidden within opaque libraries (e.g., Netty, JGroups, Java Message Service, or the Java Remote Method Invocation mechanism) to the point where a programmer rarely resorts to using low-level mechanisms like sockets. Similarly, the problem of keeping concurrent execution correct should be so hidden away under an abstraction, rather than expecting the programmer to do it manually using primitives like locks, semaphores, or barriers and avoid their various pitfalls like deadlocks, data races, or priority inversion.

Enter: distributed transactional memory (DTM) [4, 13, 6, 17, 18, 25]. DTM is an approach modeled on database transactions and transactional memory known from (single-host) multiprocessor systems [11, 19]. In this approach, the programmer simply annotates which areas of code are transactions. The DTM system transparently ensures that code executed within transactions accesses shared remote resources in a safe and consistent manner.

The thing that most starkly differentiates DTM from its database predecessors is that apart from executing read and write operations on shared resources (aka shared objects), these can provide interfaces for other or different operations. Classically, this can be an operation like increment, which both reads and writes the state of a shared object atomically, or stack operations like push and pop. These can also be more complex, computation-intensive, programmer-specified operations that execute just about any code, and can include code with side effects. By extension (and in contrast to non-distributed TM) it matters which processor or which network node executes this code. Therefore, it makes a difference what model of execution an implementation uses. The data flow (DF) model entails shared objects being migrated to the client that uses them (while maintaining only a single copy of the object in the system). In such a case the computation and side effects will be performed on whatever host the object is migrated to for executing an operation. In the control flow (CF) model, shared objects are bound to individual hosts and do not migrate, so the execution of their operations is performed always on the object’s ”home” host. Both models have their advantages and disadvantages, but a unique feature of CF is that it allows to delegate computation to remote hosts. This allows client transactions to ”borrow” computational power from remote resource servers. In effect shared resources can act as both shared memory and web services. This provides greater flexibility in designing and implementing distributed systems.

However, to the best of the authors’ knowledge, there are currently no well-performing CF DTM systems, and certainly none that would match a top notch DF system like HyFlow2 [25] which implements the Transaction Forwarding Algorithm (TFA) [13] for concurrency control. Its predecessor, HyFlow [17] implements additional algorithms, including DTL2 (a distributed variant of TL2 [7]) and has CF support. However, as the authors show, HyFlow2 outperforms HyFlow by a significant margin, to the point of obsoleting it. The authors suppose, however, that there is no intrinsic problem stopping a CF system from being equally efficient as a DF system.

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In our previous research we introduced Atomic RMI [24], a CR DTM system that implements the *Supremum Versioning Algorithm* (SVA), which employs pessimistic concurrency control. Typically, TM systems employ the optimistic approach, where, generally speaking, a client transaction executes regardless of other transactions running in parallel and performs validation only when it finishes executing (at commit time). If two transactions try to access the same shared object, and one of them writes to it, they conflict and one of the transactions aborts and restarts (in an optimized TM, this occurs as soon as possible, to waste least work). When a transaction aborts, it should not change the system state, so aborting transactions must revert the objects they modified. Alternatively, they work on local copies and merge them with the original object only on a successful commit. Pessimistic concurrency control is a different approach, known in database transactions (e.g., two-phased locking) and brought to TM in [14, 3] and our earlier work [27, 28]. It involves transactions waiting until they have permission to access shared objects. In effect, potentially conflicting operations are postponed until they no longer conflict. Thus transactions, for the most part, avoid forced aborts, and thus, they also naturally avoid the problems stemming from irrevocable operations. The advantage of using pessimistic concurrency control over the optimistic approach in a CF DTM system is that it does not require as stringent assumptions with respect to what code can be part of transactions or operations executed on shared resources. Optistic concurrency control assumes implicitly that local operations executed within a transaction (i.e., those not executed on shared objects directly) do not need to be compensated for in the event of an abort, and can be safely re-executed. This is not the case for a class of irrevocable operations, which include acquiring and releasing locks, system I/O, sending and consuming network messages. Such operations are more likely to occur as the complexity of the code of transactions increases, as can potentially happen with applications using CF to delegate complex tasks to remote resources. The problem was mitigated in non-distributed TM by using special irrevocable transactions that preempt other transactions and cannot abort (but only run one-at-a-time) [29], or providing multiple versions of objects that transactions view for reads [2, 10]. In other cases, irrevocable operations are simply forbidden in transactions (e.g., in Haskell [10]). Meanwhile, since pessimistic TM does not depend on aborts to maintain consistency, the problem of irrevocable operations is mitigated for free.

The perceived problem of pessimistic TMs is that they struggle with performance (see e.g., [14]), but the main feature of SVA is the use of early release to confer a performance improvement. Specifically, it uses *a priori* knowledge to detect when a transaction will certainly perform no further operations on a particular shared object, and, in such a case, allows other transactions to then access that object, rather than having them wait until the first transaction commits. However, since SVA is agnostic with respect to whether operations on shared objects modify state or are read-only, it often perceives potential conflicts, where those conflicts will not occur *de facto*, and thus limits parallelism more than strictly necessary. For this reason, Atomic RMI performs similarly to HyFlow (with DTL2) and therefore is significantly outperformed by HyFlow2 (with TFA).

In our most recent research [30], we introduced a new pessimistic TM concurrency control algorithm called *Optimized SVA* (OptSVA), which introduces a number of optimizations to the basic *modus operandi* of its predecessor, mainly: heavy use of buffering rather than in-place modifications, read operation parallelization, early release of shared objects on last write instead of last operation of any kind, and asynchrony which allows transactions to delegate some tasks that require waiting to separate thread and proceed with other computation in the mean time. These optimizations reduce the amount of scenarios where one transaction has to wait for another and in this way improve the transactional throughput.

However, OptSVA operates in a system model more typical for non-distributed TMs, where shared objects are simple variables that can be either read from or written to and cannot be trivially lifted to a general CF model. Then, for instance, if the first operation on a variable in the transaction is a write, an OptSVA transaction does not need to synchronize the state of the variable with other transactions, but instead can simply store the value it writes to a buffer and proceed. Furthermore, all subsequent reads are local—they only depend on the value written by the current transaction, so there is no need to synchronize before executing those reads either. On the other hand, the CF model allows operations to execute more complex code on the shared object, and it may be that, for example, a write only modifies some field **a** of the object, but a subsequent read accesses its field **b**. In that case OptSVA would need to synchronize the state of the object with other transactions before the read operation is allowed to execute, but it does not supply a mechanism for that, so it is not directly useful in the CF model.

The first contribution of the paper is to lift the OptSVA algorithm to the CF system model, by re-designing it with complex programmer-defined shared objects in mind. In this way we introduce OptSVA-CF, a more general variant of OptSVA, that nevertheless is similarly capable of highly parallel execution of concurrent transactions and that shares its predecessor’s strong safety guarantees.

The second contribution of this paper is Atomic RMI 2, an implementation of OptSVA-CF that builds on Atomic RMI. As with Atomic RMI, Atomic RMI 2 provides a simple-to-use API that allows programmers to implement consistent transactions as simply as if using much simpler mechanisms, such as distributed coarse-grained locking. However, the application of OptSVA-CF allows for the execution of these transactions to be highly parallelized. We demonstrate this in a comprehensive evaluation of Atomic RMI 2, showing that it produces a significant efficiency increase over its predecessor as well as various lock-based distributed concurrency control solutions. In addition, we show that Atomic RMI 2 performs better than, or comparably to HyFlow2 (depending on contention and operation length), but does so while avoiding aborting transactions altogether, thus allowing the use of irrevocable operations.

The paper is structured as follows. In Section 2, we
introduce the OptSVA-CF and explain how it works and how it differs from its non-CF predecessor. Then, in Section 2 we proceed to discuss the implementation details of OptSVA-CF within the framework of Atomic RMI 2, including the transactional API and the architecture. In Section 3 we evaluate Atomic RMI 2 comparing it against Atomic RMI—its predecessor, HyFlow2—a state-of-the-art DF DTM, and a number of typical lock-based distributed synchronization solutions. Then, in Section 4 we discuss related work, and conclude in Section 5.

2 OptSVA-CF

OptSVA-CF is a new pessimistic DTM concurrency control algorithm operating in the CF model. In this section we first describe the mechanisms the algorithm employs. We follow by providing a summary of OptSVA-CF.

2.1 Concurrency Control through Versioning

In versioning algorithms, when each transaction starts, it is assigned a private version for each object it plans on accessing throughout its lifetime. We denote transaction $T_i$’s private version for shared object $obj_x$ as $pv_i(obj_x)$. The private versions are assigned from a sequence of consecutive positive integers in such a way that:

a) no two transactions have the same private version for any shared object,

b) if transaction $T_i$ started before $T_j$ and they both access $obj_x$, then $pv_i(obj_x) < pv_j(obj_x)$,

c) given two transactions $T_i$ and $T_j$, if $pv_i(obj_x) < pv_j(obj_y)$ then for any shared object $obj_y$ that both transactions plan to access, $pv_i(obj_y) < pv_j(obj_y)$, and

d) if $T_i$ started before $T_j$ and no other transaction started between the two, and both plan to access $obj_x$, then they have consecutive private versions for $obj_x$, i.e., $pv_i(obj_x) = pv_j(obj_x) - 1$.

The versioning mechanism uses private versions to maintain order when accessing shared objects via objects’ local versions. That is, each shared object $obj_x$ has its own local version counter, denoted $lv(obj_x)$, which is always equal to the private version of such transaction $T_j$ that most recently finished using the object: $T_j$ committed, aborted, or determined that it would no longer need the object and released it early (see below). When $T_j$ does any of those things, it writes its own private version to the local version counter. Once $T_j$ releases the object in such a way it follows that some other transaction can safely start calling methods on the object. We determine which transaction gets to access the object next by simply selecting the transaction with the next consecutive private version, i.e., such $T_i$ whose $pv_i(obj_y) - 1 = pv_j(obj_x)$. Thus, invariably $obj_x$ can be accessed by such $T_i$, for which $pv_i(obj_x) - 1 = lv(obj_x)$, and no other transaction. Hence, if some transaction $T_i$ wants to access $obj_x$, then it may do so if $pv_i(obj_x) - 1 = lv(obj_x)$. We call this condition the access condition. On the other hand, if $T_i$ wants to access $obj_x$ and the access condition is not satisfied, then it waits until it is satisfied.

An example of how this mechanism works is shown in Fig. 1. The diagrams depict histories consisting of operations executed by transactions on a time axis. Every line depicts the operations executed by a particular transaction. The symbol $\rightarrow$ denotes a complete operation execution. The inscriptions above operation executions denote operations executed by the transactions, e.g., $obj_x.op() \rightarrow v$ denotes that an operation on variable $obj_x$ is executed by the transaction and returns some value $v$, and $tryC_i \rightarrow C_i$ indicates that transaction $T_i$ attempts to commit and succeeds because it returns $C_i$. On the other hand, the symbol $\leftarrow$ denotes an operation execution split into the invocation and the response event to indicate waiting, or that the execution takes a long time. In that case the inscription above is split between the events, e.g., a read operation execution would show $obj_x.op() \rightarrow \text{above the invocation}$, and $\rightarrow v$ over the response. If waiting is involved, the arrow $\leftarrow$ is used to emphasize a happens-before relation between two events. Annotations below events emphasize the state of counters or actions performed within the concurrency control algorithm (used as necessary).

In Fig. 1, $T_1$ and $T_2$ attempt to access shared variable $obj_y$ at the same time. Transaction $T_1$ starts first, so $pv_1(obj_y) = 1$, and $T_2$ starts second, so $pv_2(obj_y) = 2$. Since initially $lv(obj_y) = 0$, $T_2$ is not able to pass the access condition and execute an operation on $obj_y$ when it tries to, so it waits. On the other hand, $T_1$ can pass the access condition $pv_1(obj_y) - 1 = lv(obj_y)$ and it executes an operation on $obj_y$ without waiting. Once $T_1$ commits, it sets $lv(obj_y)$ to 1, so $T_1$ then becomes capable of passing the access condition and finishing executing its operation on $obj_y$. In the mean time, transaction $T_k$ can proceed to access $obj_y$ completely in parallel.

2.2 Early Release

Versioning algorithms use an early release mechanism to execute concurrent transactions partially in parallel. A transaction can release any shared object early at any point during its lifetime. This indicates that the transaction no longer has any need to use the object, so some other transaction may start using it instead. Early release can be effected automatically by the versioning algorithm. In order to release a shared object automatically, a transaction must have a priori knowledge of suprema: it must know at most how many times it will attempt to access each ob-

![Figure 1: Concurrency control via versioning.](image-url)
commit condition, and otherwise it must wait. In effect, if some transaction of the transaction that either committed or aborted last.

version counter accessing objects. Each shared object has a local terminal state. We can imagine a situation where a transaction also introduce a possibility of transactions reading inconsistent state. We can imagine a situation where a transaction

T commits before T. It means that the algorithm has a chance to forcibly abort a transaction.

Any transaction that aborts writes, marks each object in its access set as an invalid instance, and reverts each object’s state. Each transaction checks whether the object is valid before accessing it or executing commit- or abort-related activities on it. If the object is invalid, then the transaction is forced to abort instead of executing whatever operation it was supposed to execute. Since the commit operations are ordered according to private versions, a transaction will not abort unless it is impossible for any of the objects it accessed to be invalidated.

An example of how this mechanism affects execution is shown in Fig. 3. Here, T, T both access obj, and they respectively get the values of pv for obj of 1 and 2. Transaction T accesses obj first and releases it early, setting lv(obj) to 1. This allows T to pass the access condition and access obj. Transaction T subsequently attempts to commit. However, in order to commit T must pass the termination condition pv(obj) − 1 ≤ lv(obj), which will not be satisfied until T sets ltv(obj) to its own pv(obj). Hence T can only complete to commit until T commits or aborts. When T eventually aborts it invalidates obj. When T attempts to commit, it cannot do so, because obj is invalid, so it finally aborts.

2.4 Irrevocable Operations

The concurrency control algorithm is pessimistic in nature, and does not need to abort any transaction to ensure correct execution. Even so, a way to manually abort transactions is provided to the programmer, since this makes it easier to cancel a transaction mid-execution without having to manually scrub its effects. More importantly, it makes the implementation of such features as fault tolerance possible. However, in conjunction with early release, aborts also introduce a possibility of transactions reading inconsistent state. We can imagine a situation where a transaction releases an object early, and another transaction starts accessing it, but the first transaction eventually aborts. In that case the second transaction cannot be allowed to commit, since it accessed data that was later invalidated.

In versioning algorithms that allow aborts the scenario is resolved by maintaining the order in which transactions execute a commit or abort on each variable, by analogy to accessing objects. Each shared object has a local terminal version counter lv(obj) which holds the private version of the transaction that either committed or aborted last.

Irrevocable transactions are those that allow aborts but do not provide a way to manually abort them. In practice, this means that aborts are not cascaded in OptSVA-CF. However, in other transactional memory systems, cascading aborts can occur as a result of a cascade, by replacing all access condition checks with termination condition checks. This means that irrevoca-

Figure 2: Early release via upper bounds.

Figure 3: Cascading abort via versioning.
ble transactions never "accept" objects released early. The drawback is that such transactions may wait longer to access shared objects, but in return they never forcibly abort.

2.5 Operations in the Complex Object Model

Initially, versioning algorithms were conceived as operation-type agnostic, which made them suitable for use with complex shared objects using in the CF model. These objects have arbitrary interfaces, whose operations (methods) execute arbitrary computations on state that can be composed of multiple discrete variables (fields). Such operations may not be limited to reading or writing the state of the object, but may do both, or neither, or cause side-effects in the process. Furthermore, each object may have a different interface. It is therefore practical to treat such objects as black boxes with respect to their state and operations they execute.

Contrast this to simple objects, variables, used commonly in TM (see e.g., [9]), where each object has a single read operation that reads the state of the object, and a single write operation which supersedes the previous state of the object with a new state. Both operations are simple, completely transparent, and contain no side effects, which allows to better orchestrate their execution.

An apt example is that of operation locality. According to [9], a local read is a read that is preceded during transaction execution by a write on the same shared variable—because it only depends on state written by that write operation, so it does not depend on what other transactions write. A local write is a write that is followed by a write on the same variable—because whatever that first write writes is superseded by the value written by the second. Local operation executions do not impact the system outside of their transaction. Thus, buffering can be used to make them invisible to the outside world. A local write modifies a transaction-local buffer, rather than the actual object. This means that local writes de facto do not operate on shared objects, so they do not need to pass the access condition to be executed. As we showed with OptSVA [30], using such optimizations with this model means that transactions execute more in parallel, and produce tighter schedules as a result, which improves system throughput.

On the other hand, the simplification of complex objects to variables restricts the flexibility of such a system model and limits its applicability in distributed systems. This is especially true in the CF model, where a complex object can be used not only to store and retrieve data, but also to delegate more involved, possibly long-running computations to a remote host. Once the arbitrary nature of interfaces and operation semantics is removed, the latter is lost and the system model loses its expressiveness. Hence we introduce the distinction between read and write operations in the complex object model with arbitrary interfaces by requiring that each operation be classified as one of the following:

a) a read operation is any operation that executes any code (including code with side effects) and may read the shared object's state and return a value, but during execution the state is never modified,

b) a write operation is any operation that executes any code and may modify the state of the shared object, but the state is not read, nor modified,

c) an update operation is any operation that executes any code and may both modify and read the object's state and return a value.

This classification allows us to mimic the optimizations used with variable-like objects within complex object synchronization, but without knowing the details of each operation's semantics. We introduce the update operation, because we expect a typical operation on a complex object to modify its state based on the existing state of the object, hence to behave both like a read and write. However, such an operation is difficult to make invisible and parallelize. On the other hand, "pure" writes, can be expected to be rare, but they do not need to view the state to execute, so more optimizations apply to them. Specifically, they can also be made to execute on an "empty" buffer without prior synchronization. Thus, we keep them apart from updates. Note that the complex shared object may still contain composite state, consisting of some number of independent variables, and read, write, and update operations are not required to read and/or modify the state holistically. Whether or not a particular operation will only read state written locally or whether it requires synchronization depends largely on how objects buffering is implemented.

2.6 Buffering

When creating buffers for variable-like objects, given the semantics of the two available operations, it is simply a matter of copying a value from a shared variable to some local variable. Such a buffer can also be locally written to without knowledge of its state, since the new value of the variable supersedes the old. Thus, local writes can simply write to uninitialized local variables.

Given the composite state of complex objects and arbitrary semantics of operations, two types of buffers are needed. The first, a copy buffer, is one that copies the entire state of a shared object, and can be used to both locally read and modify the object. Such a buffer can be used to read a released object or restore an object during abort. However, since the state of the original object is copied, in order to create a copy buffer the transaction must check fulfill the access condition before doing so. Such a copy buffer is not universal, since it cannot be used to execute local writes without prior synchronization.

Thus, we introduce a second buffer type. A log buffer is an object that maintains the interface of the original shared object but none of its state. When a method is executed on the object, the buffer logs the method and its parameters. The method may be executed completely, assuming that it does not need any state other than local data to do so. In that case, any changes the method does to the state are tracked and stored. If this is impossible, the method will not execute, apart from being logged. The log buffer can be
applied to the original object to update the state of the latter. If some method was pre-executed before applying the log, its effects are applied to the state of the original object. If a method was not previously executed, it is executed on the original object at the time the log is being applied. Given the log buffer does not use the object’s state, it can be used to execute write operations without prior synchronization. Since write operations modify the object’s state without viewing it, write operations are always capable of executing methods on the log buffer in place, and do not need to commute the execution to the point when the buffer is applied.

Since the CF semantics require that computations are performed wherever the shared object is located in the distributed system, either type of buffer resides on the same host in a distributed system as the original object. Otherwise, not only would the assumptions of the CF model be violated, but if the execution of operations caused any side effects, the side effects would be removed from the location of the original node.

2.7 Asynchronous Buffering

A special case occurs if a transaction only ever executes read operations on a shared object (although it may execute writes and updates on other objects). We will refer to such an object as a read-only object with respect to this transaction. In the case of such an object, synchronization needs to be done when the first read is executed, but all subsequent reads only need to use the buffer to execute. Hence, once the buffer is created, the reads execute as if they were local and the read-only object can be released.

Other transactions benefit from the object being released as soon as possible, and it is possible for a read-only object to be released even before the first read operation occurs. I.e., first write does not need to access the actual object either, as long as the state of the object is buffered. The only condition that must be satisfied to store the state of some object in the buffer is that it must pass the access condition, but otherwise it can be done at any point in the transaction. Hence, OptSVA-CF transactions attempt to buffer a read-only object as soon as they retrieve private versions at start. But, since waiting at the access condition may block the execution of operations that precede the first read on the read-only object in the code of the transaction, the buffering procedure is executed asynchronously: the transaction delegates it to a separate thread and proceeds to execute other operations as normal. The separate thread waits until the access condition for the object is met, following which the thread buffers, and immediately releases the object. Then, all reads, including the first read, execute the operation on the buffer. In effect, early release of read-only objects is potentially expedited.

We show an example of this in Fig. 4. Here, transaction $T_j$ treats $\text{obj}_j$ as a read-only object, and tries buffering it as soon as it starts. This is done in a separate thread (indicated by the line below) which waits until the access condition is satisfied for $\text{obj}_j$. Once $T_j$ releases $\text{obj}_j$, the thread immediately buffers the object and releases it. This allows transaction $T_k$ to begin accessing and modifying $\text{obj}_j$, even while $T_j$ executes two read operations in parallel using its buffer. If the buffer were not used, $T_j$ would delay operations performed by $T_k$.

Similar asynchrony is used in the case of a final modification of an object. When a transaction executes its last write or update operation on some shared object, the object is immediately buffered afterward and released. This allows all following read operations to only use the buffer, and therefore be invisible to outside transactions. The final update can only be executed if the access condition is passed, since it may need to view the state of the object to execute. However, a write may execute using the log buffer instead and without synchronization, since it does not view the state. Then, in the specific case of a write operation that is the only write operation on an object, or in case of a write operation preceded only by other write operations on that object, the transaction may not have attempted to satisfy the access condition yet. In such a case, the final write can be split into a write that executes using the log buffer without synchronization, and a procedure that subsequently updates the state of the actual object. This procedure can only be executed if the access condition is passed, but it can release the object immediately after it finishes updating the object’s state. The procedure is executed asynchronously with respect to the main body of the transaction, since it has no impact on following operations—all following operations on this object, if any, will be reads, and will read from the local buffer. In this way, the last write avoids blocking the entire transaction to wait for the access condition. In addition, the point at which the transaction must wait for the access condition for this object can be delayed to any point between the last write and the end of the transaction.

We illustrate this optimization further in Fig. 5. Here, transaction $T_i$ can pass access condition for $\text{obj}_j$ first and execute a read and a write on $\text{obj}_j$. Nevertheless $T_j$ performs operations on $\text{obj}_j$ simultaneously. First, $T_j$ executes a write, and can do so without waiting at the access condition, since it works on the log buffer rather than directly on $\text{obj}$. Meanwhile $T_i$ can execute operations on the actual object. Then $T_j$ executes another write operation, using the log buffer. Since this is the last operation execution on $\text{obj}$ in $T_j$ (which the transaction knows because $\text{wub}_j(\text{obj}_j) = \text{wc}_j(\text{obj}_j)$), $T_j$ may write the changes from the log buffer into the object. Hence, a separate thread starts at the end of the write operation and it starts waiting at the access condition. When the access condition is
2.8.1 Start

When an OptSVA-CF transaction \(T_i\) starts, it acquires a private version for each shared object in its access set. If any of these objects are read-only, the transaction also starts separate threads that clone the objects into copy buffers \(buf_i(obj_x)\) and release them afterwards.

2.8.2 Read

Whenever transaction \(T_i\) attempts to execute a read operation on some object \(obj_x\), its behavior primarily depends on whether the object is read-only, and whether it was released or not. If the object is read-only, the read operation waits until the separate read-only thread finishes buffering the object, and executes the read operation on the buffer.

Otherwise, if the object was not previously accessed, then the transaction checks if there were preceding reads or updates. If not, the transaction must wait until the access condition to \(obj_x\) is satisfied and makes a checkpoint by copying the state of the shared object to buffer \(st_i(obj_x)\).

Figure 5: Asynchronous release on last write.

2.8 Summary

OptSVA-CF operates on the basis of the versioning mechanism, using private, local, and local terminal version counters to ensure that accesses to objects and commits are performed in the order defined by private versions. The individual operations are handled as follows:

2.8.1 Start

When an OptSVA-CF transaction \(T_i\) starts, it acquires a private version for each shared object in its access set. If any of these objects are read-only, the transaction also starts separate threads that clone the objects into copy buffers \(buf_i(obj_x)\) and release them afterwards.

2.8.3 Write

In the case of an update operation, the transaction checks whether and updates were executed on the same object before. If that is the case, the transaction waits until the access condition is satisfied and makes a checkpoint by copying the state of the shared object to buffer \(st_i(obj_x)\). In addition, the transaction will also apply the log buffer \(log_i(obj_x)\) to \(obj_x\). If the transaction was previously released, the operation executes on \(obj_x\), and if the object was not read-only, the transaction applies the log buffer to \(obj_x\).

If the thread updates the state of \(obj_x\), using the copy buffer, and does so in parallel to \(T_i\). If the thread updates the state of \(obj_x\), using the copy buffer, and does so in parallel to \(T_i\).

2.8.4 Write

Pure write operations are executed in one of two ways, depending on whether the object was read or update operations executed prior. This is because updates and reads
both wait on the access condition, meaning that then the object can be operated on directly. Otherwise, the write can be performed using a log buffer. Specifically, if there were no preceding reads or updates, the transaction simply executes the operation on the log buffer. If this is the final write and there will also not be update operations on this object in the transaction, the transaction then starts a thread, which will wait at the access condition and subsequently: make a checkpoint to $st_i(\text{obj}_j)$, apply the log buffer $\log_i(\text{obj}_j)$ to the original object $\text{obj}_i$, copy the modified object to the copy buffer $\text{buf}_i(\text{obj}_j)$, and release $\text{obj}_j$. Meanwhile, the transaction’s main thread proceeds.

If there were preceding reads or updates, the transaction operates using the up-to-date object that is already under its control. Making a checkpoint would be redundant, but the transaction checks whether any objects were invalidated, and if so, aborts. Otherwise, it executes the code of the operation on the object, and if this was the last write or update operation on $\text{obj}_j$, then $\text{obj}_j$ is cloned to $st_i(\text{obj}_j)$ and released. The last step is not done in a separate thread, since the transaction already has access to $\text{obj}_j$.

2.8.5 Commit

When the transaction commits it waits for extant threads to finish in the case such threads are still running for read-only objects and objects that are being released after last write. Afterwards, the transaction waits until the commit condition is satisfied for all objects in its access set. Then, if the transaction did not access a particular object at any time, it makes a checkpoint. If it only ever executed writes on an object, the transaction applies the log buffer to the object. If the object was not released, the transaction releases it. Afterward, the transaction checks whether any object was invalidated, and aborts if that is the case. Otherwise, the transaction updates the local terminal versions of all objects and finishes execution. No further actions may be performed by the transaction after the commit finishes executing.

2.8.6 Abort

When the transaction aborts, just like with commit, it waits for the appropriate threads to finish, and for the commit condition to be satisfied. Then, each object in the transaction’s access set is restored from $st_i(\text{obj}_j)$, unless some other transaction that previously aborted already restored it to an older version beforehand. Then, the transaction updates the local terminal versions of all objects and finishes execution. No further actions may be performed by the transaction after the abort finishes executing.

2.9 Consequences of Model Generalization

OptSVA-CF is based on the optimizations introduced in OptSVA, but applies them to a different, more universal system model. The complex object model is more general, since a variable-like object can be implemented as a reference cell, a complex object with one field, a read operation returning its value, and a write operation setting the old value to a new one.

Given such a specification, OptSVA-CF will execute the same way as OptSVA with one exception. Given a transaction that executes a write operation on some object $\text{obj}_x$ followed by a read operation on $\text{obj}_y$, OptSVA-CF will execute the write without synchronization, but must synchronize before the read executes. This is because the changes in the log buffer must be applied to the actual object and copy buffer before the read proceeds. Hence, the read might be forced to wait until the access condition for $\text{obj}_y$ is satisfied. In contrast, OptSVA will allow the read to proceed without synchronization, since it is recognizable as a local operation, and therefore completely dependent on the preceding write.

In effect, given reference cells, there are certain executions that will be allowed by OptSVA that are not allowed by OptSVA-CF. Since these OptSVA executions are tighter than their equivalent executions in OptSVA-CF, OptSVA admits a higher level of parallelism. Therefore, OptSVA-CF trades generality for performance.

2.10 Properties

We briefly demonstrate the safety, liveness, and progress guarantees of OptSVA-CF.

2.10.1 Safety

In [20] we demonstrate that OptSVA is last-use opaque. The requirements in synchronizing complex objects mean that if OptSVA-CF is used to model the system model of OptSVA, OptSVA-CF allows a subset of executions allowed by OptSVA. Hence, OptSVA-CF is also last-use opaque [21][22]. This implies that OptSVA-CF is serializable, recoverable, preserves real-time order, and does not allow overwriting a value once an object is released.

Furthermore, if the manual abort operation is never used within a given system, OptSVA-CF never causes any transaction to abort, meaning that such OptSVA-CF executions are indistinguishable from opaque [9] (as shown in [23], and by analogy to [3]). This extends the guarantees above to imply that a transaction never reads from an aborted transaction, and cascading aborts are avoided.

2.10.2 Liveness

There are two types of occurrence where an operation can wait. The first is waiting on an access condition, or the similar condition when a transaction attempts to commit or abort. In this case, the condition is satisfied in the order enforced by transactions’ private versions for specific objects. Since private versions are consecutive integers and since they are acquired atomically by the transaction, it is impossible for a circular wait to occur. The other case of waiting is during transaction start, when private versions are acquired. In order for this to be done atomically, transactions lock a series of locks before getting private versions, and release the locks afterwards. These locks are always acquired in accordance to an arbitrary global order, regardless of transaction. That eliminates the possibility
that a circular wait occurs during start. Since circular wait
cannot occur among transactions, OptSVA-CF is deadlock
free.

2.10.3 Progress

Any transaction in OptSVA-CF can either abort manually
or forcibly. In order for a transaction \( T_i \) to abort forcibly, there must be some transaction \( T_j \) that forces \( T_i \) to abort, i.e., such \( T_j \) that accessed some object \( obj_x \) and released it
before \( T_i \) accessed \( obj_x \), and \( T_j \) must have aborted after \( T_i \)
accessed \( obj_x \). Thus for every forcibly aborted transaction,
there must be another aborted transaction. Hence, given
any set of conflicting transactions, there will be at least one
transaction that will not be forcibly aborted (but it will
be manually aborted). Therefore, OptSVA-CF is strongly
progressive [9].

3 Architecture

Atomic RMI 2 is a framework that supplies transactional
concurrency control in a distributed system on top of Java
RMI. The framework uses OptSVA-CF for synchronization.

We give an overview of the system architecture in Fig. 6
with two client nodes and two server nodes, showing the
flow of control when remote methods are executed on
shared objects. In general, the system may contain any
number of independent client and server nodes. Each of
the server nodes hosts a number of discrete, uniquely iden-
tifiable shared objects, whose methods are called by client
nodes as operations. Any node can act simultaneously as
client as well as a server. Each shared object is located at
exactly one specific node (as opposed to the object being
copied or moved to other nodes, or being replicated on sev-
eral nodes) and all operations invoked on that object will
execute on the node that hosts it.

Shared objects can be accessed remotely from client
nodes in the system by calling methods specified by the
object’s interface (as per the control flow model). A shared
objects interface is defined by the programmer and consists
of a number of methods that can be called on the object.

Clients execute operations on shared objects as part of
transactions. We show an example of a transaction defini-
tion in Fig. 8. The programmer declares a transaction using
the API provided by Atomic RMI 2 (Fig. 8), by creating a
Transaction object, which is responsible for starting and
stopping transactional execution. A transaction can be de-
defined as irrevocable at this point, meaning it will never be
forced to abort, because it will not access objects that are
released early, instead waiting for the preceding transaction
to commit or abort. Then, that object is used to declare

```
interface Transaction {
    Transaction(boolean irrevocable);
    <T> T updates(T obj);
    <T> T writes(T obj);
    <T> T reads(T obj);
    <T> T accesses(T obj);
    <T> T updates(T obj, int maxUpdates);
    <T> T writes(T obj, int maxWrites);
    <T> T reads(T obj, int maxReads);
    <T> T accesses(T obj, int maxRd, int maxWr, int maxUpd);
    void start(Transactional runnable);
    void commit();
    void retry();
    void abort();
}
interface Transactional {
    void atomic(Transaction t);
}
```

Figure 8: Atomic RMI 2 transaction interface.

The semantics of the operations are defined by the pro-
grammer and can be anything from simple gets and sets,
to complex methods executing arbitrary server-side code,
accessing a database, or even invoking other remote ob-
jects. Each method must be annotated as either a read
operation, a write or an update operation. We give an ex-
ample of a shared object interface for a bank account in
Fig. 7.

```
interface Account extends Remote {
    @Access(Mode.READ) int balance();
    @Access(Mode.UPDATE) void deposit(int value);
    @Access(Mode.WRITE) void withdraw(int value);
    @Access(Mode.WRITE) void reset();
}
```

Figure 7: Shared object interface example (a bank ac-
count).

```
interface Transaction {
    Transaction(boolean irrevocable);
    <T> T updates(T obj);
    <T> T writes(T obj);
    <T> T reads(T obj);
    <T> T accesses(T obj);
    <T> T updates(T obj, int maxUpdates);
    <T> T writes(T obj, int maxWrites);
    <T> T reads(T obj, int maxReads);
    <T> T accesses(T obj, int maxRd, int maxWr, int maxUpd);
    void start(Transactional runnable);
    void commit();
    void retry();
    void abort();
}
interface Transactional {
    void atomic(Transaction t);
}
```

```
Transaction t = new Transaction(irrevocable=false);
Account a = t.accesses(registry.locate("A"), 1, 0, 1);
Account b = t.updates(registry.locate("B"), 1);
t.start(new Transactional() {
    void atomic(Transaction t) {
        a.deposit(100);
        b.deposit(100);
        if (a.balance() < 0)
            t.abort();
    }
});
```

Figure 9: Transaction definition example (with manual
abort).
the transaction’s preamble, where the programmer specifies which objects will be used by the transaction and how, by passing the reference retrieved from the RMI registry to method reads, updates, writes, or accesses—the latter if more than one kind of operation may be executed on the object. The programmer can use variants these methods to also provide suprema for any object used by the transaction. The suprema indicate the maximum number of times the transaction will execute read-only, write, and update methods on each shared object throughout the execution of the code. In the example in Fig. 9 the preamble declares the transaction will invoke at most one read-only, at most one update method on A, and at most one update method on shared object B.

In practice, suprema do not have to be derived manually, but instead static analysis [20] or the type system [27] can be used. If suprema are not given, infinity is assumed (and the system maintains guarantees). If suprema are provided though, the underlying concurrency control algorithm uses them to effect early release, and in this way increase the level of parallelism between concurrent transactions.

### 3.1 Instrumentation

When accesses are declared within the preamble, an object stub is created. This stub is then used within the code of the transaction to invoke methods on the shared objects, as with ordinary RMI stubs. The difference between an ordinary RMI stub and an Atomic RMI 2 stub is that the latter does not forward method calls to the shared object directly, but instead uses a proxy object. Proxy objects are created dynamically on the node hosting the shared object in question at the same time as the stub is created by the transaction. Each proxy object links one specific shared object on the server side with one specific transaction (object) on the client side. Proxy objects implement the interfaces of the shared objects they are linking, and their role is to inject the concurrency control code of OptSVA-CF before and after the invocation of specific methods of the shared object. The injection is done via reflection, which supplies the necessary flexibility, allowing arbitrary methods to easily be supplemented with concurrency control. In theory, proxy objects could be located either on the server side or the client side, but since the communication between the proxy and the shared object is much more frequent than that between the transaction and the proxy, placing them on the server-side incurs lower overheads. In addition, if the proxy is placed on the server, it can easily manage copy and log buffers, which must be placed on the server to preserve the CF model—methods executed using buffers should have side effects on the same node as the original object. Proxy objects can be decommissioned once a transaction that created them finishes executing.

### 3.2 Transactional Code

Once the preamble of the transaction is in place, the transaction’s code can be specified. This is done by creating an object implementing the Transactional interface, whose atomic method then defines the logic of the transaction. In general, a transaction can contain virtually any code between its start and either commit or abort. This specifically means that apart from operations on shared objects, any local operations, e.g., irrevocable operations, can be present within. As an example of a simple transaction, in Fig. 9 the programmer specifies a transaction that transfers 100 currency from one bank account to another. Thus, an anonymous Transactional object is created and within the atomic method, the withdraw method is called on object A (the stub for shared object A), after which the deposit method is called on B (the stub for B). The programmer can rest assured the concurrency control algorithm will synchronize the execution of this code so that no other transaction in the system interferes with the execution in a way that would violate its consistency. If the transaction reaches the end of its code it attempts to commit. The programmer is also given the option to abort or retry the entire transaction manually by using the transaction object and invoking either the abort or retry method. Here, the transaction is rolled back if it turns out that the balance on account A fell below 0 as a result of executing withdraw.

Since Atomic RMI 2 implements a pessimistic concurrency control algorithm, transactions never abort as a result of conflict (see discussion in Section 2). It is possible for a transaction to abort as a result of explicitly invoking abort, which can cause a cascading rollback, or if a failure is detected (discussed below). Any transaction can be prevented from ever aborting though by specifying it as irrevocable.

### 3.3 Executor Thread

OptSVA-CF calls for asynchronous execution using separate threads. Given the cost of overhead that starting a thread creates, Atomic RMI 2 uses one executor thread per JVM. The executor thread is always running and transactions assign it tasks. Each task consists of a condition and code. The code of the task is meant to be executed only when the condition is satisfied. Once the thread receives a task, it checks whether it can be immediately executed. If not, it queues up the task and waits until any of the two counters that can impact the condition change value (1v and 1tv). When any of the counters change, the thread re-evaluates the relevant conditions and executes the task, if the condition so allows.

### 3.4 Fault Tolerance Mechanisms

In distributed environments faults are a fact of life, so any DTM system must have mechanisms to deal with them. Atomic RMI 2 handles two basic types of failures: remote object failures and transaction failures.

Remote object failures are straightforward and the responsibility for detecting them and alarming Atomic RMI 2 falls onto the mechanisms built into Java RMI. Whenever a remote object is called from a transaction and it cannot be reached, it is assumed that this object has suffered a failure and an exception is thrown. The programmer may then choose to handle the exception by, for example, re-
running the transaction, or compensating for the failure. Remote object failures follow a crash-stop model of failure: any object that crashed is removed from the system.

On the other hand, a client performing some transaction can crash causing a transaction failure. Such failures can occur before a transaction releases all its objects and thus make them inaccessible to all other transactions. The objects can also end up in an inconsistent state. For these reasons transaction failures need also to be detected and mitigated. Atomic RMI 2 does this by having remote objects check whether a transaction is responding. If a transaction fails to respond to a particular remote object (times out), it is considered to have crashed, and the object performs a rollback on itself: it reverts its state and releases itself. If the transaction actually crashed, all of its objects will eventually do this and the state will become consistent. On the other hand, if the crash was illusory and the transaction tries to resume operation after some of its objects rolled themselves back, the transaction will be forced to abort when it communicates with one of these objects.

4 Evaluation

In this section we present the results of a practical evaluation of Atomic RMI 2 in the context of other distributed TM concurrency control mechanisms operating in a similar system model.

4.1 Frameworks

The first framework we use for evaluation is Atomic RMI [24], a distributed pessimistic CF TM implementing SVA [27] on top of Java RMI. SVA uses the bare supremum versioning mechanism described in Section 2.1 and is operation-type agnostic. The comparison against SVA shows off the optimizations introduced in OptSVA-CF, especially since the two algorithms are implemented using the same technology and give the same guarantees. The second framework we compare Atomic RMI 2 against is HyFlow2 [25], a state-of-the-art distributed TM system implemented in Scala. HyFlow2 implements the optimistic Transactional Forwarding Algorithm (TFA) [17] and operates in the data flow model. TFA is opaque but does not have provision for irrevocable operations.

We also compare all three TM systems against distributed concurrency control solutions based on locks. Specifically, we use distributed mutual exclusion locks (marked Mutex) and read-write locks (R/W Locks), both custom-tailored and implemented on top of Java RMI. In both solutions a lock is created for every shared object in the system. Each locking solution has two variants. The first variant is a straightforward usage where every transaction locks every object from its access set when it commences, and releases each of object on commit. This is equivalent to a conservative (strong) strict two-phase locking solution and satisfies opacity. We denote this variant as S2PL. The second variant represents non-strict two-phase locking (2PL), and is a more advanced implementation from the programmer’s point of view. Here, each transaction also initially locks each of the objects in its access set, but the programmer determines the last access on each object in the transaction’s access set and manually releases the lock early (prior to commit). Non-strict two-phase locking satisfies last-use opacity under the assumption that last access is always determined correctly. We denote the second variant as 2PL. Finally, we also use a solution with a single global mutual exclusion lock (GLock) that is acquired by each transaction for the duration of the transaction’s entire execution. This produces a completely sequential execution and acts as a baseline for the purpose of the comparison.

4.2 Benchmark

We perform our evaluation using a 16-node cluster connected by a 1Gb network. Each node had two quad-core Intel Xeon L3260 processors at 2.83 GHz with 4 GB of RAM each and runs OpenSUSE 13.1 (kernel 3.11.10, x86_64 architecture). We use Groovy version 2.3.8 with the 64-bit Java HotSpot(TM) JVM version 1.8 (build 1.8.0.25-b17).

The evaluation is performed using our own distributed implementation of Eigenbench [12]. Eigenbench is a flexible, powerful, and lightweight benchmark that can be used for comprehensive evaluation of multicore TM systems by simulating a variety of transactional application characteristics.

Eigenbench uses three arrays of shared objects, each of which is accessed with a different level of contention. The hot array contains some number \( n \) of objects that can be accesses by transaction in any thread. The access to objects in the hot array is controlled by the TM. The mild array contains \( n \) objects per thread. The access to these objects is also controlled by the TM, but the objects are partitioned in such a way, that no two transactions ever conflict on them. The third, cold array is populated and partitioned like the mild array, but it is only accessed non-transactionally. Each object within any of the three arrays is a reference cell, i.e., an object that holds a single value, that can be either read or written to. These arrays are accessed by client transactions. Each transaction accesses semi-randomly selected objects in all three arrays in random order, with the exception that the number of accesses to each type of array is specified, and the ratio of read operations to write operations on each type of array is also specified. The benchmark has a specified locality, which is a probability with which transactions will access the same object several times. When an object is being selected by a transaction, a random number is generated, and if it is below the locality probability, the object is selected at random from the transaction’s history of objects accessed thus far. Otherwise, the object is selected randomly from the pool of all shared objects. Locality and the length of the history are parameters of the benchmark.

4.3 Results and Discussion

Fig. [10] illustrates the change of throughput (measured in the number of executed operations on shared data per second) as the number of clients increases from 64 (4 per node)
Figure 10: Throughput vs client count.

(a) 90% reads, 10% writes.
(b) 50% reads, 50% writes.
(c) 10% reads, 90% writes.

Figure 11: Throughput vs node count (hot array accesses).

(a) 90% reads, 10% writes, 5 arrays.
(b) 50% reads, 50% writes, 5 arrays.
(c) 10% reads, 90% writes, 5 arrays.
(d) 90% reads, 10% writes, 10 arrays.
(e) 50% reads, 50% writes, 10 arrays.
(f) 10% reads, 90% writes, 10 arrays.

Figure 12: Throughput vs node count (hot and mild array accesses).

(a) 90% reads, 10% writes.
(b) 50% reads, 50% writes.
(c) 10% reads, 90% writes.
to 1024 (64 per node). We show three scenarios, each executed on 16 nodes, with 10 arrays of each type per node. Each client executes 10 consecutive transactions, each with 10 operations on the hot array per transaction, with a $9 \div 1$, $5 \div 5$, or $1 \div 9$ read-to-write operation ratio. Each operation takes around 3ms to execute, not counting the overhead from synchronization, network communication, or serialization overhead. This means operations are fairly long, which represents the complex computations. The locality of operations is set to 50% with a history of 5 operations.

The graphs show that all frameworks’ throughput falls as the number of clients, and therefore contention, increases. The decline is steep until 256 clients, and it levels out by 1024 clients. All systems significantly outperform the serial execution forced through GLock. In the 90% read scenario HyFlow2 and Atomic RMI 2 outperform other frameworks by a significant margin of between 9 and 267% (not counting GLock), with the exception of R/W 2PL outperforming HyFlow2 at 64 clients. Atomic RMI 2 outperforms HyFlow2 initially (by 9–25%), but after 512 clients are introduced, HyFlow2 takes the lead (by 2–23%), and both frameworks throughputs eventually converge at the 1024 client mark. In the other two scenarios, all frameworks suffer a decrease in throughput, but Atomic RMI 2 remains relatively efficient, outperforming all other frameworks, including HyFlow2, by 9–359%. The difference stems from the write-oriented optimizations in Atomic RMI 2 that allow the framework to tighten the executions in the presence of larger contingents of write operations, just as much as is possible in read-dominated schedules: objects are acquired for writing as late as possible and released prior to commit. Meanwhile other frameworks typically do not optimize write operations to the same extent. Specifically, HyFlow2 does not release early on writes, and R/W 2PL cannot perform any optimizations on writes, apart from early release on last write. In addition a degradation in Atomic RMI 2’s performance is also partly explained by the need to introduce new threads to handle asynchrony, which can become a bottleneck and offset the gain from Atomic RMI 2’s optimizations if other threads are also running on the same node (like client threads here). Among the remaining frameworks, any 2PL always performs better than the apposite S2PL variant, and R/W performs better than Mutex. Atomic RMI performs on par with Mutex 2PL and significantly below Atomic RMI 2.

Fig. 11 shows a change in throughput with constant contention as new nodes are introduced. In this scenario, we vary the number of nodes from 4 to 16 with 5 or 10 arrays of each type hosted on each node (yielding lower and higher contention respectively), and 16 clients running per node. The remainder of parameters is as above. As more processors are introduced into the system, the number of transactions running in parallel increases, causing the throughput of all frameworks to increase as well.

In the 5-array scenarios in Fig. 11, the comparison shows that Atomic RMI 2 significantly and consistently outperforms Atomic RMI and all remaining frameworks, with the exception of HyFlow2. Specifically, Atomic RMI 2 achieves at least a 47% better throughput over Atomic RMI due to the introduced optimizations. The impact of read-only optimizations is visible in the 90% read scenario, where Atomic RMI 2 achieves up to a 201% advantage in throughput. Furthermore, the write optimizations give Atomic RMI 2 a performance boost of up to 167% over Atomic RMI in the 90% write scenario. In a more balanced scenario optimizations can be applied less often, leading to a slightly lower performance improvement of up to 72%. Note, that Atomic RMI’s performance does not change with respect to the differences in workloads among scenarios, since Atomic RMI is agnostic of operation types. HyFlow2 and Atomic RMI 2 perform similarly in read dominated and balanced scenarios, with HyFlow2 outperforming Atomic RMI 2 by up to 10% in a 16-node system, and Atomic RMI 2 outperforming HyFlow2 by as much in a 4 node system. The similarities in performance stem from special handling of read-only variables in both systems. However, in a write dominated scenario, Atomic RMI 2 has a 77% percent throughput advantage, which we again attribute to extensive write-oriented optimizations employed in OptSVA-CF.

The 10-array scenarios in Fig. 11 yield similar results, but here, Atomic RMI 2 manages to consistently outperform HyFlow2, as well as other evaluated frameworks. This is because transactions have more objects to randomly select from, transactions tend to contain shorter subsequences of operations on the same objects, which allows Atomic RMI 2 to release more objects earlier.

Fig. 12 shows changes in throughput as above, but with longer transactions, that perform mild array accesses in addition to hot array accesses. Hence each transaction performs 10 operations on the hot array and 10 operations on the mild array, in the same read-to-write ratios. Since accesses on mild arrays never lead to conflicts, the average contention is much lower in this scenario than the previous. Because of this, throughput increases for each framework. Atomic RMI 2 performs similarly to HyFlow2 in the balanced scenario (up to 2% reduction or 8% improvement), slightly better in the read dominated (8–19% improvement), and significantly better in the write dominated scenario (64–76%). Both HyFlow2 and Atomic RMI 2 perform significantly better than all other frameworks, including Atomic RMI. The results are similar to those in the previous scenario, but show that Atomic RMI 2’s advantage decreases in lower contention, which we attribute to the overhead introduced by the instrumentation and asynchronous execution.

The abort rates of Atomic RMI 2 and Atomic RMI remain at 0% throughout the evaluation, while 60–89% of HyFlow2 transactions abort and retry at least once due to conflicts, depending on the scenario (see Fig. 13). This means, that irrevocable operations are likely to be aborted and re-executed. On the other hand, Atomic RMI 2 manages to rival the efficiency of an optimistic TM system while bypassing problems with irrevocable operations.

Throughout we see that Atomic RMI 2 significantly outperforms Atomic RMI and other lock-based distributed concurrency control mechanisms, and performs similarly to or better than a state-of-the-art optimistic distributed TM, all without the need to use aborts and, thus, without complicating irrevocable operation executions, and while
employing the reflection-based mechanisms that allow to use CF model. We also see that Atomic RMI 2 performs best in read-dominated scenarios, but become really competitive in write-dominated scenarios, where the buffering-and asynchrony-related write-oriented optimizations make a real difference to throughput.

5 Related Work

Several distributed TM systems were proposed (see e.g., [4, 6, 13, 29]). Most of them replicate a non-distributed TM on many nodes and guarantee that replicas are consistent. Their programming model is different from our distributed transactions. Other systems extend non-distributed TMs with a communication layer, e.g., DistM [13] extends [6] with distributed coherence protocols. The others include HyFlow2 [24], a distributed TM operating in the data flow model that is covered in Section 4. HyFlow1 [17] is an earlier version of HyFlow2, implemented in Java on top of Aleph and DeuceSTM and included control flow and data flow concurrency control algorithms, including TFA [18] and DTL2 (a distributed version of TL2 [7]). HyFlow was compared with HyFlow2 in [25] and was shown to perform worse then its successor.

Distributed transactions are successfully used where requirements for strong consistency meet wide-area distribution, e.g., in Google’s Percolator [13] and Spanner [5]. Percolator supports multi-row, ACID-compliant, pessimistic database transactions that guarantee snapshot isolation. This is a much weaker guarantee than expected from TM systems. Another drawback in comparison to DTM is that writes must follow reads. Spanner provides semi-relational replicated tables with general purpose distributed transactions. It uses real-time clocks and Paxos to guarantee consistent reads. Spanner requires some a priori information about access sets and defers commitment like Atomic RMI 2, but aborts on conflict. Irrevocable operations are banned in Spanner. Spanner transactions provide snapshot isolation and external consistency (akin to real-time order), much weaker properties than considered sufficient in DTM.

A pessimistic TM (but not DTM) is proposed in [14], where read-only transactions execute in parallel, but transactions that update are synchronized using a global lock to execute one-at-a-time. This idea was improved upon in Pessimistic Lock Elision (PLE) [1], where a number of optimizations were introduced, including encounter-time synchronization, rather than commit-time. However, the authors show that sequential execution of update transactions yields a performance penalty. In contrast, the algorithm proposed here maintains a high level of parallelism regardless of updates. In particular, the entire transaction need not be read-only for a variable that is read-only to be read-optimized.

In [14] the authors propose pessimistic non-distributed TM that runs transactions sequentially (as in [29]) but allows parallel read-only transactions. Operations are synchronized by delaying writes of the write-set location (with busy waiting). This is done using version numbers of transactions. In contrast, Atomic RMI 2 uses object versions for similar purposes, which enables early release. However, direct comparison is difficult, because [14] aims at non-distributed environments with fast access, while Atomic RMI 2 assumes network communication with overheads.

SemanticTM [8] is another pessimistic (non-distributed) TM. Rather than using versioning or blocking, transactions are scheduled and place their operations in bulk into a producer-consumer queues attached to variables. The instructions are then executed by a pool of non-blocking executor threads that use statically derived access sets and dependencies between operations to ensure the right order of execution. The scheduler ensures that all operations of one transaction are executed consistently and in the right order. The transactions cannot abort, either forcibly nor by manual override, but any operations can have redundant executions (causing problems with irrevocable operations). On the other hand, SemanticTM is wait-free, whereas OptSVA-CF is only deadlock-free.

6 Conclusions

The paper introduced OptSVA-CF, a new pessimistic TM concurrency control algorithm for use with the CF model with complex objects. The algorithm is based on suprime versioning, which allows manual aborts (although does not abort transactions due to conflicts), and implements a number of optimizations based on OptSVA, aiming at better parallel execution of concurrent transactions. Generalization from variables to complex objects requires categorization of operations into three discrete groups: reads, aborts, and updates. It also requires that two types of buffers are used, a copy buffer and a log buffer, the latter of which allows pure initial writes to execute without synchronization, just like in the variable model. OptSVA-CF is implemented as Atomic RMI 2, which performs better than its predecessor: Atomic RMI with SVA, as well as a number of lock-based synchronization mechanisms. It can also outperform HyFlow2, a state-of-the-art optimistic DF DTM, and does so without aborting (i.e. without problems with irrevocable operations). Given this, we show that a pessimistic system can be as well-performing as an optimistic one, and introduce a CF DTM with competitive performance that was lacking.

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