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FACULTY OF COMPUTER SCIENCE AND ENGINEERING**



SPECIALIZED PROJECT

**STUDYING AND DEVELOPING
NONBLOCKING DISTRIBUTED MPSC QUEUES**

Major: Computer Science

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Disclaimers

I affirm that this specialized project is the product of my original research and experimentation. Any references, resources, results which this project is based on or a derivative work of have been given due citations and properly listed in the footnotes and the references section. All original contents presented are the culmination of my dedication and perserverance under the close guidance of my supervisors, Mr. Thoại Nam and Mr. Diệp Thanh Đăng, from the Faculty of Computer Science and Engineering, Ho Chi Minh City University of Technology. I take full responsibility for the accuracy and authenticity of this document. Any misinformation, copyright infrigment or plagiarism shall be faced with serious punishment.

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Chapter I Introduction

This chapter details the motivation for our research topic: “Studying and developing nonblocking distributed MPSC queues”, based on which we come to the formulation of our research question. Based on this question, we set out the objectives and the scope of our research. We end this chapter with an overview of the thesis and the structure of the rest of this document.

1.1 Motivation

The demand for computation power has always been increasing relentlessly. Increasingly complex computation problems arise and accordingly more computation power is required to solve them. Much engineering efforts have been put forth towards obtaining more computation power. A popular topic in this regard is distributed computing: The combined power of clusters of commodity hardware can surpass that of a single powerful machine. To fully take advantage of the potential of distributed computing, specialized distributed algorithms and data structures need to be devised. Hence, there exists a variety of programming environments and frameworks that directly support the execution and development of distributed algorithms and data structures, one of which is the Message Passing Interface (MPI).

Traditionally, distributed algorithms and data structures use the usual Send/Receive message passing interface to communicate and synchronize between cluster nodes. Meanwhile, in the shared memory literature, atomic instructions are the preferred methods for communication and synchronization. This is due to the historical differences between the architectural support and programming models utilized in these two areas. For a class of problems known as regular applications, the use of the traditional Send/Receive interface suffices. However, this interface poses a challenge for irregular applications (Section 2.1). Fortunately, since the introduction of specialized networking hardware such as RDMA and the improved support of the remote memory access (RMA) programming model in MPI-3, this challenge has been alleviated: irregular applications can now be expressed more conveniently with an API that’s similar to atomic operations in shared memory programming. This also implies that shared-memory algorithms and data structures can also be ported to distributed environments in a more straightforward manner. Since the design and development of shared-memory algorithms and data structures have been extensively studied, this has opened up a lot new research such as [2] on applying the principles of the shared memory literature to distributed computing.

Concurrent multi-producer single-consumer (MPSC) queue is one of those data structures that have seen many applications in shared-memory environments and plays the central role in many programming patterns, such as the actor model and the fan-out/fan-in pattern, as shown in Figure 1.

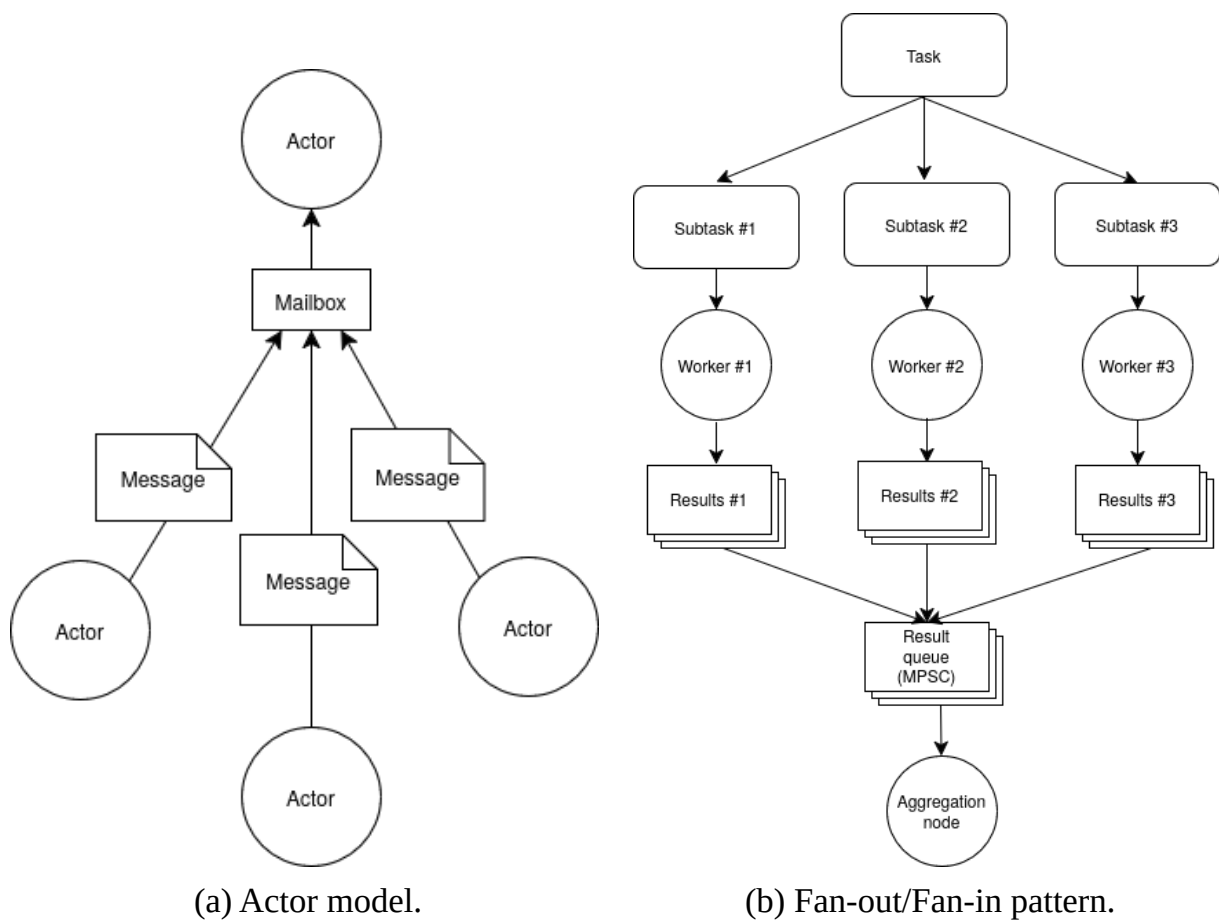


Figure 1: Some programming patterns involving the MPSC queue data structure.

In the actor model, each process or compute node is represented as an actor. Each actor has a mailbox, which exhibits MPSC queue property: Other actors can send messages to the mailbox and the owner actor extracts messages and performs computation based on these messages. The fan-out/fan-in pattern involves splitting a task into multiple subtasks to workers, then the workers queue back the result to the master, who dequeues out the results to perform further processing, such as aggregation. These patterns can be potentially useful if they can be expressed efficiently in distributed environments. However, we have found discussions of distributed MPSC queue algorithms in the current literature to be very scarce and scattered and as far as we know, none has focused on designing an efficient distributed MPSC queue. The closest we found is the Berkeley Container Library (BCL) [3] that provides many distributed data structures including a multi-producer single-consumer (MPMC) queue and multi-producer/multi-consumer (MP/MC) queue. This presents an inhibition to programmers that want to either directly use the distributed MPSC queues or express programming patterns that inherently express MPSC queue behaviors, they either have to work around the requirement or remodel their problems in another way. If a distributed MPSC queue is also provided as part of a library, this can in turn encourage many distributed applications and programming patterns that utilize the MPSC queues.

A desirable distributed MPSC queue algorithms should possess two favorable characteristics (1) scalability, the ability of an algorithm to utilize the highly concurrent nature of distributed clusters (2) fault-tolerance, the ability of an algorithm to continue running

despite the failure of some compute nodes. Scalability is important for any concurrent algorithms, as one would never want to add more compute nodes just for performance to drop. Fault-tolerance, on the other hand, is especially more important in distributed computing, as failures can happen more frequently, such as network failures, node failures, etc. Fault-tolerance is concerned with a class of properties arisen in concurrent algorithms known as progress guarantee (Section 2.4). Specifically, lock-freedom is one such property that allows an algorithm to keep running even when there's some suspended processes.

Lock-free MPSC queues and other FIFO variants, such as multi-producer multi-consumer (MPMC), concurrent single-producer single-consumer (SPSC), have been heavily studied in the shared memory literature, dating back from the 1980s-1990s [4], [5], [6] and more recently [7], [8]. It comes as no surprise that lock-free algorithms in this domain are highly developed and optimized for performance and scalability. However, most research about distributed algorithms and data structures in general completely disregard the available state-of-the-art algorithms in the shared memory literature. Because shared-memory algorithms can now be straightforwardly ported to distributed context using this programming model, this presents an opportunity to make use of the highly accumulated research in the shared memory literature, which if adapted and mapped properly to the distributed context, may produce comparable results to algorithms exclusively devised within the distributed computing domain. Therefore, we decide to take this novel route to developing new non-blocking MPSC queue algorithms: Utilizing shared-memory programming techniques, adapting potential lock-free shared-memory MSPCs to design fault-tolerant and performant distributed MPSC queue algorithms. If this approach proves to be effective, a huge intellectual reuse of the shared-memory literature into the distributed domain is possible. Consequently, there may be no need to develop distributed MPSC queue algorithms from the ground up.

1.2 Research question

“How to utilize shared-memory programming principles to model and design distributed MPSC queue algorithms in a fault-tolerant and performant manner?”

We further decompose this question into smaller subquestions:

1. Which factor contributes to the fault-tolerance and performance of a distributed MPSC queue algorithms?
2. Which shared-memory programming principle is relevant in modeling and designing distributed MPSC queue algorithms in a fault-tolerant and performant manner?
3. Which shared-memory programming principle needs to be modified to more effectively model and design distributed MPSC queue algorithms in a fault-tolerant and performant manner?

1.3 Objective

Based on what we have listed out in the previous section, we aim to:

- Investigate the principles underpinning the design of fault-tolerant and performant shared-memory algorithms.
- Investigate state-of-the-art shared-memory MPSC queue algorithms as case studies to support our design of distributed MPSC queue algorithms.
- Investigate existing distributed FIFO algorithms that can be adapted for MPSC use cases to serve as a comparison baseline.
- Model and design distributed MPSC queue algorithms using techniques from the shared-memory literature.
- Utilize the shared-memory programming model to evaluate various theoretical aspects of distributed MPSC queue algorithms: correctness and progress guarantee.
- Model the theoretical performance of distributed MPSC queue algorithms that are designed using techniques from the shared-memory literature.
- Collect empirical results on distributed MPSC queue algorithms and discuss important factors that affect these results.

1.4 Scope

The following narrows down what we're going to investigate in the shared-memory literature and which theoretical and empirical aspects we're interested in our distributed algorithms:

- Regarding the investigation of the design principles in the shared-memory literature, we focus on fault-tolerant and performant concurrent algorithm design using atomic operations and common problems that often arise in this area, namely, ABA problem and safe memory reclamation problem.
- Regarding the investigation of shared-memory MPSC queues currently in the literature, we focus on linearizable MPSC queues that support at least lock-free enqueue and dequeue operations.
- Regarding correctness, we're concerned ourselves with the linearizability correctness condition.
- Regarding fault-tolerance, we're concerned ourselves with the concept of progress guarantee, that is, the ability of the system to continue to make forward process despite the failure of one or more components of the system.
- Regarding algorithm prototyping, benchmarking and optimizations, we assume an MPI-3 setting.
- Regarding empirical results, we focus on performance-related metrics, e.g. throughput and latency.

1.5 Thesis overview

An overview of this thesis is given in Image 1.

This thesis explores the shared-memory programming model to design fault-tolerant and performant concurrent algorithms using atomic operations. Traditionally, in this aspect, two notorious problems often arise: ABA problem and safe memory reclamation. We investigate the traditional techniques used in the shared-memory literature to resolve

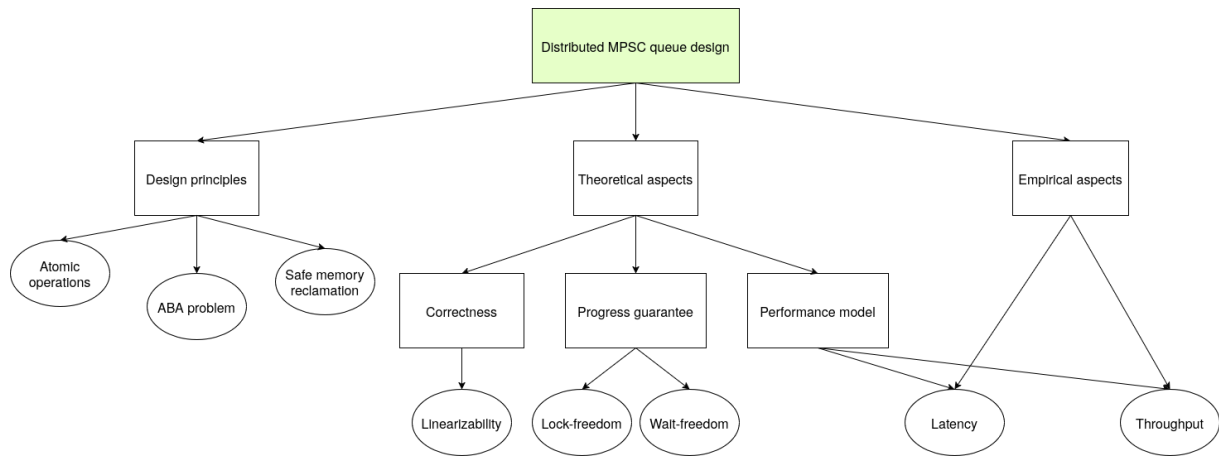


Image 1: An overview of this thesis.

these problems and appropriately adapt them to solve similar issues when designing fault-tolerant and performant distributed MPSC queues.

This thesis contributes two new distributed wait-free distributed MPSC queue algorithms. Theoretically, we're concerned ourselves with their correctness (linearizability), progress guarantee (lock-freedom and wait-freedom) which has an implication on their fault-tolerance and their performance model, specifically, the latency and throughput models as they scale to multiple nodes.

This thesis concludes with an empirical analysis of our novel algorithms to see if their actual behavior matches our theoretical performance model, interpret these results and discuss its implication.

1.6 Structure

The rest of this report is structured as follows:

Chapter II discusses the theoretical foundation this thesis is based on. As mentioned, this thesis investigates the principles of shared-memory programming and the existing state-of-the-art shared-memory MPSC queues. We then explore the utilities offered by MPI-3 to implement distributed algorithms modeled by shared-memory programming techniques.

Chapter III surveys the shared-memory literature for state-of-the-art queue algorithms, specifically MPSC queues. We specifically focus on non-blocking shared-memory algorithms that have the potential to be adapted efficiently for distributed environment. This chapter additionally surveys existing distributed FIFO algorithms to serve as a comparison baseline for our novel distributed MPSC queue algorithms.

Chapter IV introduces our novel distributed MPSC queue algorithms, designed using shared-memory programming techniques and inspired by the selected shared-memory MPSC queue algorithms surveyed in Chapter III. It specifically presents our adaptation efforts of existing algorithms in the shared-memory literature to make their distributed implementations feasible and efficient. This chapter also introduces existing FIFO

queues in the literature adapted for MPSC use cases. We aim to keep the adaptation as least intrusive as possible for fairness.

Chapter V discusses various interesting theoretical aspects of our distributed MPSC queue algorithms in Chapter IV, specifically correctness (linearizability), progress guarantee (lock-freedom and wait-freedom), performance model. Our analysis of performance model helps back our empirical findings in Chapter VI, together, they work hand-in-hand to help us discover optimization opportunities.

Chapter VI details our benchmarking metrics and elaborates our benchmarking setup. We aim to demonstrate some preliminary results on how well our novel MPSC queue algorithms, additionally compared to existing distributed FIFO queues. Finally, we discuss important factors that affect the runtime properties distributed MPSC queue algorithm, which have partly been explained by our theoretical analysis in Chapter V.

Chapter VII concludes what we have accomplished in this thesis and considers future possible improvements to our research.

Chapter II Background

2.1 Irregular applications

Irregular applications are a class of programs particularly interesting in distributed computing. They are characterized by:

- Unpredictable memory access: Before the program is actually run, we cannot know which data it will need to access. We can only know that at run time.
- Data-dependent control flow: The decision of what to do next (such as which data to access next) is highly dependent on the values of the data already accessed. Hence the unpredictable memory access property because we cannot statically analyze the program to know which data it will access. The control flow is inherently engraved in the data, which is not known until runtime.

Irregular applications are interesting because they demand special techniques to achieve high performance. One specific challenge is that this type of applications is hard to model in traditional MPI APIs using the Send/Receive interface. This is specifically because using this interface requires a programmer to have already anticipated communication within pairs of processes before runtime, which is difficult with irregular applications. The introduction of MPI remote memory access (RMA) in MPI-2 and its improvement in MPI-3 has significantly improved MPI's capability to express irregular applications comfortably. This will be explained further in Section 2.7.

2.1.1 Actor model as an irregular application



Figure 2: Actor model visualization.

Actor model in actuality is a type of irregular application supported by the concurrent MPSC queue data structure.

Each actor can be a process or a compute node in the cluster, carrying out a specific responsibility in the system. From time to time, there's a need for the actors to communicate with each other. For this purpose, the actor model offers a mailbox local to each actor. This mailbox exhibits MPSC queue behaviors: Other actors can send messages to the mailbox to notify the owner actor and the owner actor at their leisure repeatedly reads extracted message from its mailbox. The actor model provide a simple programming model for concurrent processing.

The reasons why the actor model being an irregular application are straightforward to see:

- Unpredictable memory access: The cases in which one actor can anticipate which one of the other actors can send it a message are pretty rare and application-specific. As a general framework, in an actor model, the usual assumption is that any number of actors can try to communicate with an actor at some arbitrary time. By this nature, the communication pattern is unpredictable.
- Data-dependent control-flow: If an actor A sends a message to another actor B, and when B reads this message, B decides to send another message to another actor C. As we can see, the control-flow is highly engraved in the messages, or in other words, the messages drive the program flow, which can only be known at runtime.

2.1.2 Fan-out/Fan-in pattern as an irregular application

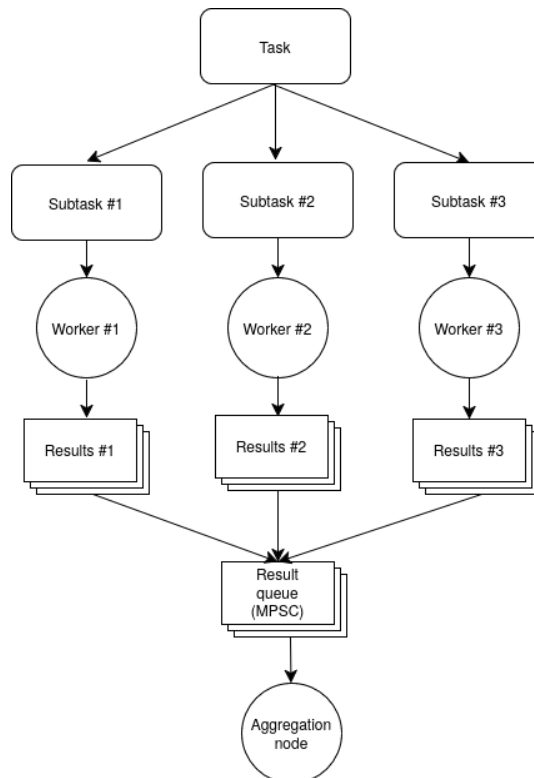


Figure 3: Fan-out/Fan-in pattern visualization.

The fan-out/fan-in pattern is another type of irregular application supported by the concurrent MPSC queue data structure.

In this pattern, there's a big task that can be splitted into subtasks to be executed concurrently on some work nodes. In the execution process, each worker produces a result set, each enqueued back to a result queue located on an aggregation node. The aggregation node can then dequeue from this result queue to perform further processing. Clearly, this result queue exhibits MPSC behavior.

The fan-out/fan-in pattern exhibits less irregularity than the actor model, however. Usually, the worker nodes and the aggregation node are known in advance. The aggregation node can anticipate Send calls from the worker nodes. Still, there's a degree of irregularity that this pattern exhibit: How can the aggregation node know how many Send calls a worker nodes will issue? This is highly driven by the task and the data involved in this task, hence, we have the data-dependent control-flow property. One can still statically calculate or predict how many Send calls a worker node will issue, however, this is problem-specific. Therefore, the memory access pattern is somewhat unpredictable. Notice that if supported by a concurrent MPSC queue data structure, the fan-out/fan-in pattern is free from this burden of organizing the right amount of Send/Receive calls. Thus, combining with the MPSC queue, the fan-out/fan-in pattern becomes more general and easier to program.

We have seen the role the MPSC queue plays in supporting irregular applications. It's important to understand what really comprises an MPSC queue data structure.

2.2 MPSC queue

Multiple-producer, single-consumer (MPSC) queue is a specialized concurrent first-in first-out (FIFO) data structure. A FIFO is a container data structure where items can be inserted into or taken out of, with the constraint that the items that are inserted earlier are taken out of earlier. Hence, it's also known as the queue data structure. The process that performs item insertion into the FIFO is called the producer and the process that performs items deletion (and retrieval) is called the consumer.

In concurrent queues, multiple producers and consumers can run concurrently. One class of concurrent FIFOs is the MPSC queue, where one consumer may run in parallel with multiple producers.

The reasons we're interested in MPSC queues instead of the more general multiple-producer, multiple-consumer (MPMC) queue data structures are that (1) high-performance and high-scalability MPSC queues are much simpler to design than MPMCs while (2) MPSC queues are powerful enough to solve certain problems, as demonstrated in Section 2.1. The MPSC queue in actuality is an irregular application in and out of itself:

- Unpredictable memory access: As a general data structure, the MPSC queue allows any process to be a producer or a consumer. By nature, its memory access pattern is unpredictable.
- Data-dependent control-flow: The consumer's behavior is entirely dependent on whether and which data is available in the MPSC queue. The execution paths of MPSC queues can vary, based on the queue contention i.e. some processes may

backoff or retry some failed operations, this scenario often arise in lock-free data structures.

As an implication, some irregular applications can actually “push” the “irregularity burden” to the distributed MPSC queue, which is already designed for high-performance and fault tolerance. This provides a comfortable level of abstraction for programmers that need to deal with irregular applications.

2.3 Correctness condition of concurrent algorithms

Correctness of concurrent algorithms is hard to defined, regarding the semantics of concurrent data structures like MPSC queues. One effort to formalize the correctness of concurrent data structures is the definition of **linearizability**. A method call on the FIFO can be visualized as an interval spanning two points in time. The starting point is called the **invocation event** and the ending point is called the **response event**. **Linearizability** informally states that each method call should appear to take effect instantaneously at some moment between its invocation event and response event [9]. The moment the method call takes effect is termed the **linearization point**. Specifically, suppose the followings:

- We have n concurrent method calls m_1, m_2, \dots, m_n .
- Each method call m_i starts with the **invocation event** happening at timestamp s_i and ends with the **response event** happening at timestamp e_i . We have $s_i < e_i$ for all $1 \leq i \leq n$.
- Each method call m_i has the **linearization point** happening at timestamp l_i , so that $s_i \leq l_i \leq e_i$.

Then, linerizability means that if we have $l_1 < l_2 < \dots < l_n$, the effect of these n concurrent method calls m_1, m_2, \dots, m_n must be equivalent to calling m_1, m_2, \dots, m_n **sequentially**, one after the other in that order.



Figure 4: Linearization points of method 1, method 2, method 3, method 4 happens at $t_1 < t_2 < t_3 < t_4$, therefore, their effects will be observed in this order as if we call method 1, method 2, method 3, method 4 sequentially.

Linearizability is widely used as a correctness condition because of (1) its composability [10] (if every component in the system is linearizable, the whole system is linearizable), which promotes modularity and ease of proof (2) its compatibility with human intuition, i.e. linearizability respects real-time order [10]. Naturally, we choose linearizability to be the only correctness condition for our algorithms.

2.4 Progress guarantee of concurrent algorithms

Progress guarantee is a criteria that only arises in the context of concurrent algorithms. Informally, it's the degree of hinderance one process imposes on another process from completing its task. In the context of sequential algorithm, this is irrelevant because there's only ever one process. Progress guarantee has an implication on an algorithm's performance and fault-tolerance, especially in adverse situations, as we will explain in the following sections.

2.4.1 Blocking algorithms

Many concurrent algorithms are based on locks to create mutual exclusion, in which only some processes that have acquired the locks are able to act, while the others have to wait. While lock-based algorithms are simple to read, write and verify, these algorithms are said to be **blocking**: One slow process may slow down the other faster processes, for example, if the slow process successfully acquires a lock and then the operating system (OS) decides to suspend it to schedule another one, this means until the process is awoken, the other processes that contend for the lock cannot continue.

Blocking is the weakest progress guarantee one algorithm can offer, it allows one process to impose arbitrary impedance to any other processes, as shown in Figure 5.

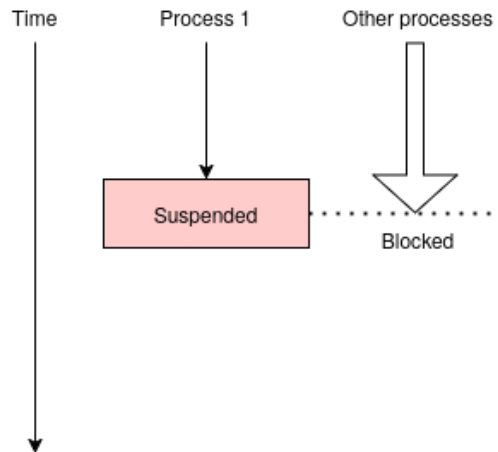


Figure 5: Blocking algorithm visualization: When a process is suspended, it can potentially block other processes from making further progress.

Blocking algorithms introduces many problems such as:

- **Deadlock:** There's a circular lock-wait dependencies among the processes, effectively prevent any processes from making progress.
- **Convoy effect:** One long process holding the lock will block other shorter processes contending for the lock.
- **Priority inversion:** A higher-priority process effectively has very low priority because it has to wait for another low priority process.

Furthermore, if a process that holds the lock dies, this will render the whole program unable to make any progress. This consideration holds even more weight in distributed computing because of a lot more failure modes, such as network failures, node failures, etc.

Therefore, while blocking algorithms, especially those using locks, are easy to write, they do not provide **progress guarantee** because **deadlock** or **livelock** can occur and its use of mutual exclusion is unnecessarily restrictive. Fortunately, there are other class of algorithms which offer stronger progress guarantees.

2.4.2 Non-blocking algorithms

An algorithm is said to be **non-blocking** if a failure or slow-down in one process cannot cause the failure or slow-down in another process. Lock-free and wait-free algorithms are two especially interesting subclasses of non-blocking algorithms. Unlike blocking algorithms, they provide stronger degrees of progress guarantees.

2.4.2.1 Lock-free algorithms

Lock-free algorithms provide the following guarantee: Even if some processes are suspended, the remaining processes are ensured to make global progress and complete in bounded time. In other words, a process cannot cause hinderance to the global progress of the program. This property is invaluable in distributed computing, one dead or suspended process will not block the whole program, providing fault-tolerance. Designing lock-

free algorithms requires careful use of atomic instructions, such as Fetch-and-add (FAA), Compare-and-swap (CAS), etc which will be explained in Section 2.5.



Figure 6: Lock-free algorithm: All the live processes together always finish in a finite amount of steps.

2.4.2.2 Wait-free algorithms

Wait-freedom offers the strongest degree of progress guarantee. It mandates that no process can cause constant hinderance to any running process. While lock-freedom ensures that at least one of the alive processes will make progress, wait-freedom guarantees that any alive processes will finish in finite number of steps. Wait-freedom can be desirable because it prevents starvation. Lock-freedom still allows the possibility of one process having to wait for another indefinitely, as long as some still makes progress.



Figure 7: Wait-free algorithm: Any live process always finishes in a finite amount of steps.

2.5 Popular atomic instructions in designing non-blocking algorithms

In non-blocking algorithms, finer-grained synchronization primitives than simple locks are required, which manifest themselves as atomic instructions. Therefore, it's necessary to get familiar with the semantics of these atomic instructions and common programming patterns associated with them.

2.5.1 Fetch-and-add (FAA)

Fetch-and-add (FAA) is a simple atomic instruction with the following semantics: It atomically increments a value at a memory location x by a and returns the previous value just before the increment. Informally, FAA's effect is equivalent to the function in Procedure 1, assuming that the function is executed atomically.

Procedure 1: `int fetch_and_add(int* x, int a)`

```
1 old_value = *x
2 *x = *x + a
3 return old_value
```

Fetch-and-add can be used to create simple distributed counters.

2.5.2 Compare-and-swap (CAS)

Compare-and-swap (CAS) is probably the most popular atomic operation instruction. The reason for its popularity is (1) CAS is a **universal atomic instruction** with the **consensus number** of ∞ , which means it's the most powerful atomic instruction [11] (2) CAS is implemented in most hardware (3) some concurrent lock-free data structures such as MPSC queues are more easily expressed using a powerful atomic instruction such as CAS.

The semantics of CAS is as follows. Given the instruction `CAS(memory location, old value, new value)`, atomically compares the value at `memory location` to see if it equals `old value`; if so, sets the value at `memory location` to `new value` and returns `true`; otherwise, leaves the value at `memory location` unchanged and returns `false`. Informally, its effect is equivalent to the function in Procedure 2.

Procedure 2: `bool compare_and_swap(int* x, int old_val, int new_val)`

```
1 if (*x == old_val)
2 |   *x = new_val
3 |   return true
4 return false
```

Compare-and-swap is very powerful and consequently, pervasive in concurrent algorithms and data structure.

Non-blocking concurrent algorithms often utilize CAS as follows. The steps 1-3 are retried until success.

1. Read the current value `old_value = read(memory location)`.
2. Compute `new_value` from `old_value` by manipulating some resources associated with `old_value` and allocating new resources for `new_value`.

3. Call CAS(memory location, old value, new value). If that succeeds, the new resources for new value remain valid because it was computed using valid resources associated with old value, which has not been modified since the last read. Otherwise, free up the resources we have allocated for new value because old value is no longer there, so its associated resources are not valid.

This scheme is, however, susceptible to the ABA problem, which will be discussed in Section 2.6.1.

2.5.3 Load-link/Store-conditional (LL/SC)

Load-link/Store-conditional is actually a pair of atomic instructions for synchronization.

Semantically, load-link returns a value currently located at a memory location x while store-conditional sets the memory location x to a value v if there's no other writes to x since the last load-link call, otherwise, the store-conditional call would fail.

Intuitively, LL/SC provides an easier synchronization primitive than CAS: LL/SC ensures that a store-conditional can only succeed if there's no access to a memory location, while CAS can still succeed in this case if the value at the memory location does not change. Due to this property, LL/SC is not vulnerable to the ABA problem (see Section 2.6.1). However, CAS is in fact as powerful as LL/SC, considering that they can implement each other [11].

Practically, store-conditional can still fail even if there's no writes to the same memory location since the last load-link call. This is called a spurious failure. For example, consider the following generic sequence of events:

1. Thread X calls load-link on x and loads out v .
2. Thread X computes a new value v' .
3. Some *exceptional event* happens (discussed below). Assume that no other threads access x during this time.
4. Thread X calls store-conditional to store v' to x . It *should succeed* but *fails* anyways.

Exceptional events that can cause the store-conditional to fail spuriously include:

- Cache line flushing: If the cache line that caches the memory location x is written back to memory, logically, the memory location x has been accessed and therefore, the store-conditional fails.
- Context switch: If thread x is swapped out by the OS, cache lines may be invalidated and flushed out, which consequently leads to the first scenario.

LL/SC even though as powerful as CAS, is not as widespread as CAS, in fact, as of MPI-3, only CAS is supported.

2.6 Common issues when designing non-blocking algorithms

2.6.1 ABA problem

ABA problem is a notorious problem associated with the compare-and-swap atomic instruction. Because CAS is so widely used in non-blocking algorithms, ABA problem almost has to always be accounted for.

As a reminder, here's how CAS is often utilized in non-blocking concurrent algorithms: The steps 1-3 are retried until success.

1. Read the current value `old value = read(memory location)`.
2. Compute new value from old value by manipulating some resources associated with old value and allocating new resources for new value.
3. Call `CAS(memory location, old value, new value)`. If that succeeds, the new resources for new value remain valid because it was computed using valid resources associated with old value, which has not been modified since the last read. Otherwise, free up the resources we have allocated for new value because old value is no longer there, so its associated resources are not valid.

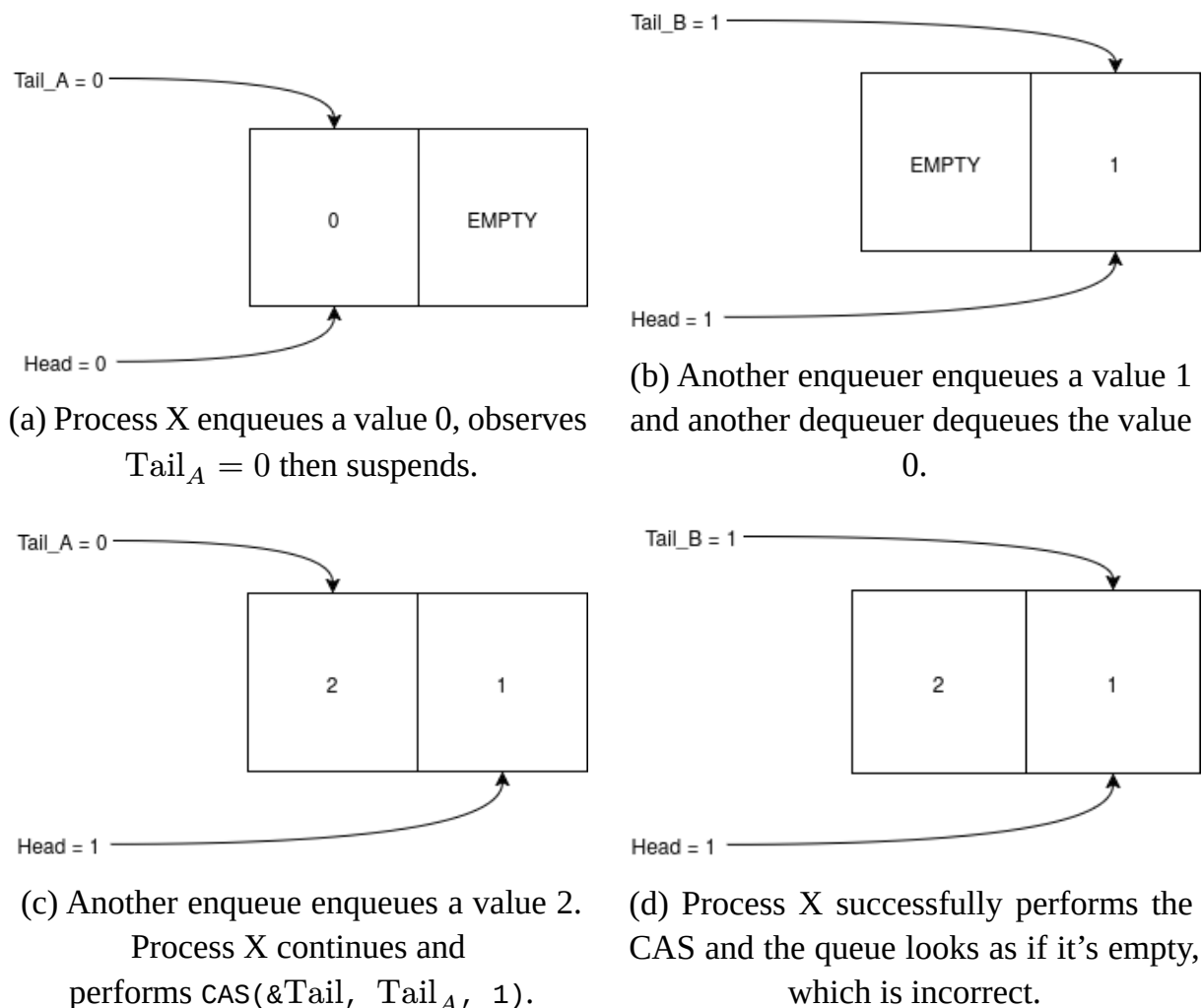


Figure 8: ABA problem in a 2-entry circular queue with Head and Tail pointers.

As hinted, this scheme is susceptible to the notorious ABA problem. The following scenario illustrate and example of ABA problem:

1. Process 1 reads the current value of memory location and reads out A.
2. Process 1 manipulates resources associated with A, and allocates resources based on these resources.
3. Process 1 suspends.
4. Process 2 reads the current value of memory location and reads out A.
5. Process 2 CAS(memory location, A, B) so that resources associated with A are no longer valid.
6. Process 3 CAS(memory location, B, A) and allocates new resources associated with A.
7. Process 1 continues and CAS(memory location, A, new value) relying on the fact that the old resources associated with A are still valid while in fact they aren't.

ABA problem arises fundamentally because most algorithms assume a memory location is not accessed if its value is unchanged.

A specific case of ABA problem is given in Figure 8.

To safe-guard against ABA problem, one must ensure that between the time a process reads out a value from a shared memory location and the time it calls CAS on that location, there's no possibility another process has CAS-ed the memory location to the same value.

2.6.2 Safe memory reclamation problem

The problem of safe memory reclamation often arises in concurrent algorithms that dynamically allocate memory. In such algorithms, dynamically-allocated memory must be freed at some point. However, there's a good chance that while a process is freeing memory, other processes contending for the same memory are keeping a reference to that memory. Therefore, deallocated memory can potentially be accessed, which is erroneous.

An example of unsafe memory reclamation is given in Figure 9.

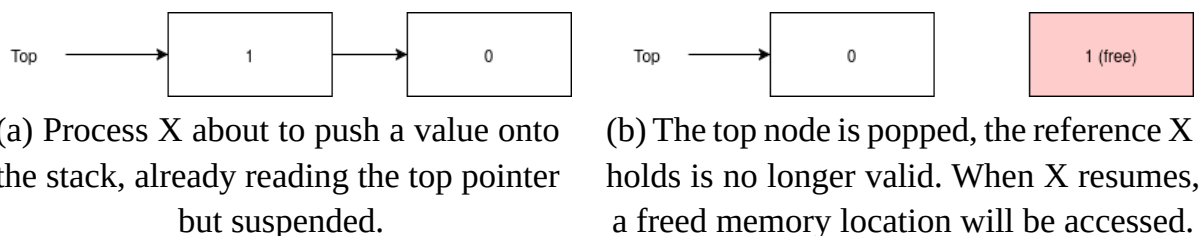


Figure 9: Unsafe memory reclamation in a LIFO stack.

Solutions to this problem must ensure that memory is only freed when no other processes are holding references to it. In garbage-collected programming environments, this problem can be conveniently push to the garbage collector. In non-garbage-collected programming environments, however, custom schemes must be utilized.

2.7 MPI-3 - A popular distributed programming library interface specification

MPI stands for message passing interface, which is a **message-passing library interface specification**. Design goals of MPI includes high availability across platforms, efficient communication, thread-safety, reliable and convenient communication interface while still allowing hardware-specific accelerated mechanisms to be exploited [1].

2.7.1 MPI-3 RMA

RMA in MPI RMA stands for remote memory access. As introduced in the first section of Section Chapter II, RMA APIs is introduced in MPI-2 and its capabilities are further extended in MPI-3 to conveniently express irregular applications. In general, RMA is intended to support applications with dynamically changing data access patterns where the data distribution is fixed or slowly changing [1]. This is very similar to the properties of irregular applications as discussed in Section 2.1. In such applications, one process, based on the data it needs, knowing the data distribution, can compute the nodes where the data is stored. However, because data access pattern is not known, each process cannot know whether any other processes will access its data. Using the traditional Send/Receive interface, both sides need to issue matching operations by distributing appropriate transfer parameters. This is not suitable, as previously explain, only the side that needs to access the data knows all the transfer parameters while the side that stores the data cannot anticipate this.

2.7.2 MPI-RMA communication operations

RMA only requires one side to specify all the transfer parameters and thus only that side to participate in data communication.

To utilize MPI RMA, each process needs to open a memory window to expose a segment of its memory to RMA communication operations such as remote writes (MPI_PUT), remote reads (MPI_GET) or remote accumulates (MPI_ACCUMULATE, MPI_GET_ACCUMULATE, MPI_FETCH_AND_OP, MPI_COMPARE_AND_SWAP) [1]. These remote communication operations only requires one side to specify.

2.7.3 MPI-RMA synchronization

Besides communication of data from the sender to the receiver, one also needs to synchronize the sender with the receiver. That is, there must be a mechanism to ensure the completion of RMA communication calls or that any remote operations have taken effect. For this purpose, MPI RMA provides **active target synchronization** and **passive target synchronization**. In this document, we're particularly interested in **passive target synchronization** as this mode of synchronization does not require the target process of an RMA operation to explicitly issue a matching synchronization call with the origin process, easing the expression of irregular applications [12].

In **passive target synchronization**, any RMA communication calls must be within a pair of MPI_Win_lock/MPI_Win_unlock or MPI_Win_lock_all/MPI_Win_unlock_all. After the unlock call, those RMA communication calls are guaranteed to have taken effect. One can also force the completion of those RMA communication calls without the need for the call to unlock using flush calls such as MPI_Win_flush or MPI_Win_flush_local.

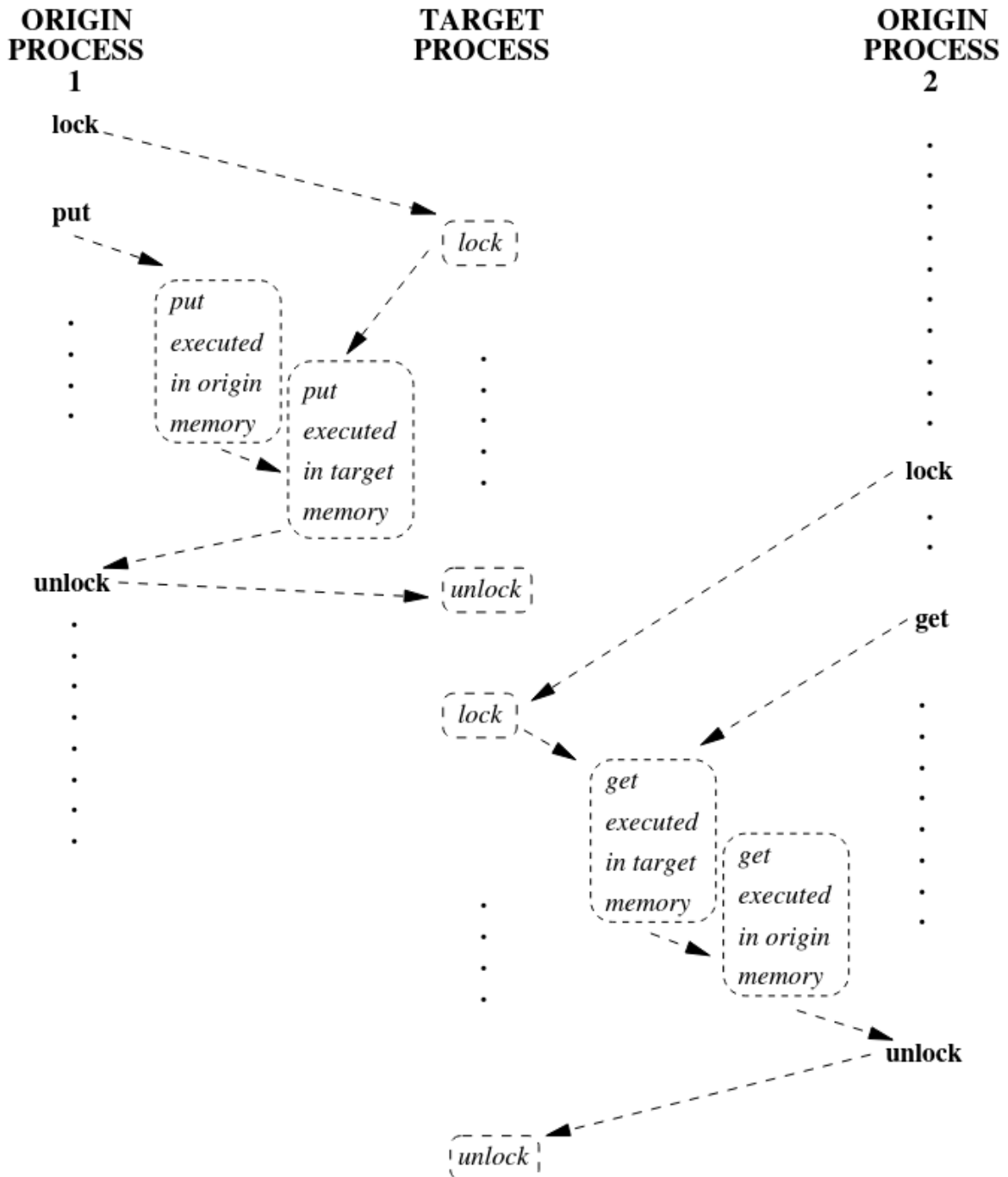


Figure 10: An illustration of passive target communication. Dashed arrows represent synchronization (source: [1]).

2.8 Pure MPI - A porting approach of shared memory algorithms to distributed algorithms

In pure MPI, we use MPI exclusively for communication and synchronization. With MPI RMA, the communication calls that we utilize are:

- Remote read: `MPI_Get`
- Remote write: `MPI_Put`
- Remote accumulation: `MPI_Accumulate`, `MPI_Get_accumulate`, `MPI_Fetch_and_op` and `MPI_Compare_and_swap`.

For lock-free synchronization, we choose to use **passive target synchronization** with `MPI_Win_lock_all`/`MPI_Win_unlock_all`.

In the MPI-3 specification [1], these functions are specified as follows:

Operation	Usage
<code>MPI_Win_lock_all</code>	Starts and RMA access epoch to all processes in a memory window, with a lock type of <code>MPI_LOCK_SHARED</code> . The calling process can access the window memory on all processes in the memory window using RMA operations. This routine is not collective.
<code>MPI_Win_unlock_all</code>	Matches with an <code>MPI_Win_lock_all</code> to unlock a window previously locked by that <code>MPI_Win_lock_all</code> .

Table 1: Specification of `MPI_Win_lock_all` and `MPI_Win_unlock_all`.

The reason we choose this is 3-fold:

- Unlike **active target synchronization**, **passive target synchronization** does not require the process whose memory is being accessed by an MPI RMA communication call to participate in. This is in line with our intention to use MPI RMA to easily model irregular applications like MPSC queues.
- Unlike **active target synchronization**, `MPI_Win_lock_all` and `MPI_Win_unlock_all` do not need to wait for a matching synchronization call in the target process, and thus, is not delayed by the target process.
- Unlike **passive target synchronization** with `MPI_Win_lock`/`MPI_Win_unlock`, multiple calls of `MPI_Win_lock_all` can succeed concurrently, so one process needing to issue MPI RMA communication calls do not block others.

An example of our pure MPI approach with `MPI_Win_lock_all`/`MPI_Win_unlock_all`, inspired by [12], is illustrated in the following:

```

MPI_Win_lock_all(0, win);

MPI_Get(...); // Remote get
MPI_Put(...); // Remote put
MPI_Accumulate(..., MPI_REPLACE, ...); // Atomic put
MPI_Get_accumulate(..., MPI_NO_OP, ...); // Atomic get
MPI_Fetch_and_op(...); // Remote fetch-and-op
MPI_Compare_and_swap(...); // Remote compare and swap
...

MPI_Win_flush(...); // Make previous RMA operations take effects
MPI_Win_flush_local(...); // Make previous RMA operations take
effects locally
...

MPI_Win_unlock_all(win);

```

Listing 3: An example snippet showcasing our synchronization approach in MPI RMA.



Figure 11: An illustration of our synchronization approach in MPI RMA.

Chapter III Related works

3.1 Non-blocking shared-memory MPSC queues

There exists numerous research into the design of lock-free shared memory MPMCs and SPSCs. Interestingly, research into lock-free MPSC queues are noticeably scarce. Although in principle, MPMC queues and SPSC queues can both be adapted for MPSC queues use cases, specialized MPSC queues can usually yield much more performance. In reality, we have only found 4 papers that are concerned with direct support of lock-free MPSC queues: LTQueue [7], DQueue [13], WRLQueue [14] and Jiffy [8]. Table 2 summarizes the characteristics of these algorithms.

MPSC queues	LTQueue [7]	DQueue [13]	WRLQueue [14]	Jiffy [8]
ABA solution	Load-link/ Store-conditional	Incorrect custom scheme (*)	Custom scheme	Custom scheme
Memory reclamation	Custom scheme	Incorrect custom scheme (*)	Custom scheme	Custom scheme
Progress guarantee of dequeue	Wait-free	Wait-free	Blocking (*)	Wait-free
Progress guarantee of enqueue	Wait-free	Wait-free	Wait-free	Wait-free
Number of elements	Unbounded	Unbounded	Unbounded	Unbounded

Table 2: Characteristic summary of existing shared memory MPSC queues. The cell marked with (*) indicates that our evaluation contradicts with the authors's claims.

3.1.1 LTQueue

To our knowledge, LTQueue [7] is the earliest paper that directly focuses on the design of a wait-free shared memory MPSC queue.

This algorithm is wait-free with $O(\log n)$ time complexity for both enqueues and dequeues, with n being the number of enqueueers due to a novel timestamp-update scheme and a tree-structure organization of timestamp.

The basic structure of LTQueue is given in Figure 12. In LTQueue, each enqueueer maintains an SPSC queue that only it and the dequeuer access. This SPSC queue must additionally support the `readFront` operation which returns the front element currently in the SPSC. The SPSC can be any implementations that conform to this interface. In the original paper, the SPSC is represented as a simple linked-list.

The rectangular nodes at the bottom in Figure 12 represents an enqueueer, whose SPSC contains items with 2 fields: value and timestamp. Every enqueueer has to timestamp

its data before enqueueing. The timestamps can be obtained using a distributed counter shared by all the enqueueers.

The purpose of timestamping is to determine the order to dequeue the items from the local SPSCs. To efficiently maintain the timestamps and determine which SPSC to dequeue from first, a tree structure with a min-heap property is built upon the enqueueer nodes. The original algorithm leaves the exact representation of the tree open, for example, the arity of the tree, which is shown to be 2 in Figure 12. The circle-shaped nodes in this figure represents the nodes in this tree structure, which are shared by all processes. Each node stores the minimum timestamp along with the owner enqueueer's rank (an identifier given to a process) in the subtree rooted at that node. After every modification to the local SPSC, i.e. after an enqueue and a dequeue, the changes must be propagated up to the root node.



Figure 12: LTQueue structure.

To dequeue, the dequeuer simply looks at the root node to determine the rank of the enqueueer to dequeue its SPSC.

The fundamental idea contributes to LTQueue's wait-freedom is the wait-free timestamp-propagation procedure. If there's a change to an enqueueer's SPSC, the timestamp of any nodes that lie on the path from the enqueueer to the root node are refreshed. The timestamp-refreshing procedure is simple:

1. Call load-link on the node's (timestamp, rank).
2. Look at all the timestamps of the node's children and determine the minimum timestamp and its owner rank.
3. Call store-conditional to store the new minimum timestamp and the new owner rank to the current node.

Notice that due to contention, the timestamp-refreshing procedure can fail. In that case, the timestamp-propagation procedure simply retries the timestamp-refreshing procedure one more time. This second call, again can fail. However, after this second call, the node's timestamp is guaranteed to be up-to-date. The intuition behind this is demonstrated in Figure 13. Furthermore, because every node is refreshed at most twice, the timestamp-refresh procedure should finish in a finite number of steps.



Figure 13: LTQueue structure.

The LTQueue algorithm avoids ABA entirely by utilizing load-link/store-conditional. This represents a challenge to directly implementing this algorithm in distributed environment.

The memory reclamation responsibility is handled by the SPSC structure, which is pretty trivial with a custom scheme.

The design of each enqueueer maintaining a separate SPSC allows multiple enqueueer to successfully enqueues its data in parallel without stepping on the others' feet. This can potentially scale well to a large number of processes. However, scalability may be limited due to potentially growing contention during timestamp propagation. The performance of LTQueue in shared-memory environments may still have a lot of room for improvement, i.e. more cache-awareness design, avoiding unnecessary contention, etc. Nevertheless, their timestamp-refreshing scheme is interesting in and out of itself and can potentially inspire the design of new algorithms. In fact, LTQueue's idea is core to one of our optimized distributed MPSC queue algorithm, Slotqueue (Section 4.4). Because DQueue does not use a proper memory reclamation scheme, ABA problem is consequently not properly resolved.

3.1.2 DQueue

DQueue [13] focuses on optimizing performance, aiming to be cache-friendly and avoid expensive atomic instructions such as CAS.

The basic structure of DQueue is demonstrated in Figure 14.



Figure 14: DQueue structure.

The global queue where data is represented as a linked list of segments. A segment is simply a contiguous array of data items. This design allows for unbounded queue capacity while still allowing a fair degree of random access within a segment. This allows us to use indices to index elements in the queue, thus permitting the use of inexpensive FAA instructions to swing the head and tail indices.

Each enqueuer maintains a local buffer to batch enqueued items before flushing to the global queue. This helps prevent contention and play nice with the cache. To enqueue an item, an enqueuer simply FAA the head index to reserve a slot in the global queue, the obtained index is stored along with the data in the local buffer so that when flushing the local buffer, the enqueuer knows where to write the data into the global queue. Note that while flushing, an index may point to a not-yet-existent slot in the global queue. Therefore, new segments must be allocated on the fly and CAS-ed to the end of the queue.

The dequeuer dequeues the items by looking at the head index. If the queue is not empty but the slot at the head index is empty, the dequeuer utilize a helping mechanism by looking at all enqueueers to help them flush out the local buffer. After this, the head slot is guaranteed to be non-empty and the dequeuer can finally dequeues out this value.

The ABA problem is solved by relying on its safe memory reclamation scheme. In DQueue, CAS is only used to update the tail pointer to point to the newly allocated segment. Therefore, ABA problem in DQueue only involves pointers. This means that if a proper memory reclamation scheme is used, ABA problem cannot occur.

DQueue relies on a dedicated garbage collection thread to reclaim segments that have been exhausted by the dequeuer. However, this should be a careful process as even though some segment has been exhausted, some enqueueers can still hold an index that references one these segments. DQueue implements this by reclaiming all exhausted

segments if there is no enqueueer holding an index referencing these segments. Unfortunately, we believe DQueue's scheme is unsafe. Specifically, as described, DQueue allows the garbage collection thread to reclaim non-adjacent segments in the global queue, without patching any of the next pointers. Any segment just before a reclaimed segment would point to a deallocated next segment. By definition, this segment was not reclaimed because it is referenced by an enqueueer. This means this enqueueer cannot traverse the next pointer chain to get to the end of the queue without accessing an already-deallocated segment.

If adapted to distributed environment, the flushing may be expensive, both from the point-of-view of the enqueueer and the dequeuer. If the dequeuer has to help every enqueueer to flush their local buffer, which should always result in at least one remote operation, the cost would be prohibitively high. Similarly, each flush requires the enqueueer to issue at least one remote operation, but this is at least acceptable as flushing is infrequent.

Still, we can still see that the pattern of maintaining a local buffer inside each enqueueer repeating throughout the literature, which we can definitely apply when designing distributed MPSC queues.

3.1.3 WRLQueue

WRLQueue [14] is a lock-free MPSC queue specifically designed for embedded real-time system. Its main purpose is to avoid excessive modification of storage space.

WRLQueue is simply a pair of buffers, one is worked on by multiple enqueueers and the other is worked on by the dequeuer. The structure of WRLQueue is shown in Figure 15.

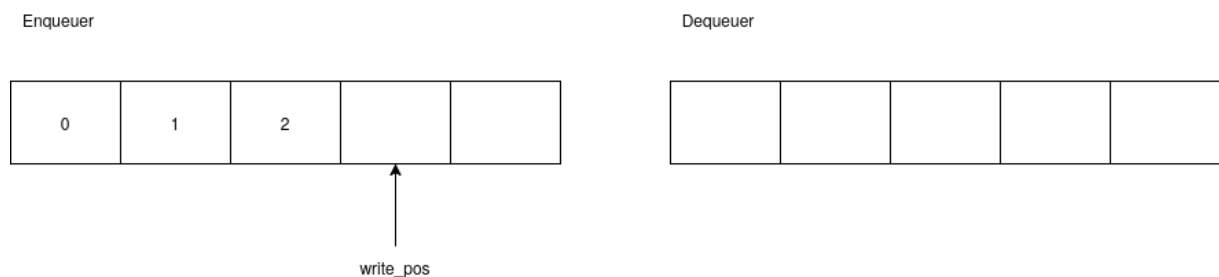


Figure 15: WRLQueue structure.

The enqueueers batch their enqueues and write multiple elements onto the buffer at once. They use the usual scheme of FFA-ing the tail index (*write_pos* in Figure 15) to reserve their slots and write data items at their leisure.

The dequeuer upon invocation will swap its buffer with the enqueueer's buffers to dequeue from it, as in Figure 16. However, WRLQueue explicitly states that the dequeuer has to wait for all enqueue operations to complete in the other buffer before swapping. If an enqueue suspends or dies, the dequeuer will experience a slow-down, this clearly violates the property of non-blocking. Therefore, we believe that WRLQueue is blocking, concerning its dequeue operation.

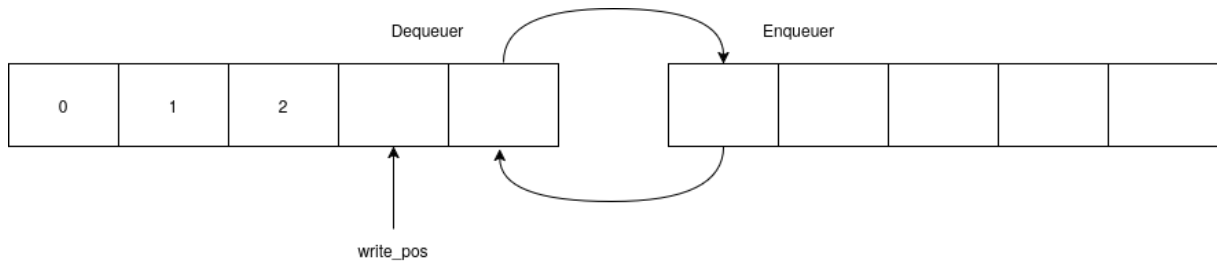


Figure 16: WRLQueue dequeue operation

3.1.4 Jiffy

Jiffy [8] is a fast and memory-efficient wait-free MPSC queue by avoiding excessive allocation of memory.

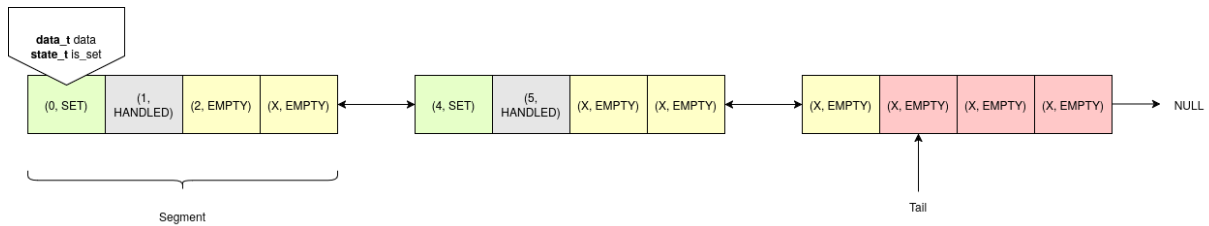


Figure 17: Jiffy structure.

Like DQueue, Jiffy represents the queue as a doubly-linked list of segments as in Figure 17. This design again allows Jiffy to be unbounded while using head and tail indices to index elements. Each segment contains a pointer to a dynamically allocated array of slots, instead of directly storing the array. Each slot in the segment contains the data item and a state of that slot (`state_t` in the figure). There are 3 states: SET, EMPTY and HANDLED. Initially, all slots are EMPTY. Instead of keeping a global head index, there are per-segment Head indices pointing to the first non-HANDLED slot. However, there is still one global Tail index shared by all the processes.

To enqueue, each enqueuer would FAA the Tail to reserve a slot. If the slot isn't in the linked list yet, it tries to allocate new segments and CAS them at the end of the linked list until the slot is available. It then traverses to the desired segment by following the previous pointers starting from the last segment. It then writes the data and sets the slot's state to SET. Notice that EMPTY slots actually have two substates. If an EMPTY slot is before the Tail index, that slot is actually reserved by an enqueuer but has not been set yet, while the EMPTY slots after the Tail index are truly empty.

To dequeue, the dequeuer would start from the Head index of the first segment, scanning until it finds the first non-HANDLED slot before the end of the queue. If there's no such slot, the queue is empty and the dequeuer would return nothing. If this slot is SET, it simply reads the data item in this slot and sets it to HANDLED. If this slot is EMPTY, that means this slot has been reserved by an enqueuer that hasn't finished. In this case, the dequeuer performs a scan forward to find the first SET slot. If not found, the dequeuer returns nothing. Otherwise, it continues to repeatedly scan all slots between the first non-HANDLED and the last found SET slot until the first SET slot between in this interval is

unchanged between 2 scans. Only then, the dequeuer would return the data item in this SET slot and mark it as HANDLED.

Similar to DQueue, CAS is only used when appending new segments at the end of the queue. Therefore, ABA problem can only occur with pointers. If a proper memory reclamation scheme is utilized.

Regarding memory reclamation, while the dequeuer is scanning the queue, it will reclaim any segments with only HANDLED slots. We can see there's potentially a pitfall similar to the one DQueue runs into here. To avoid this pitfall, Jiffy takes the following measures:

- When scanning the queue and the dequeuer sees that a segment contains only HANDLED slots, it only reclaims the dynamically-allocated array in the segment, which consumes the most memory, while still keeping the linked-list structure in tact. Therefore, if any enqueueer is holding a reference to a segment before the partially-reclaimed segment, it can still traverse the next pointer chain safely.
- To fully reclaim a segment, when partially reclaim a segment, it is added to a garbage list. Note that the first segments that contain only HANDLED slots can be fully reclaimed right when the dequeuer performs the scan. When a segment is fully reclaimed, any segment in the garbage list that precedes this segment is also fully reclaimed.

3.2 Distributed FIFO queues

FIFO queues	FastQueue [3]	dLTQueue + baseline SPSC (1)	Slotqueue + base- line SPSC (1)
Supported patterns	Multi-producer or Multi-consumer	Multi-producer Single-consumer	Multi-producer Single-consumer
ABA solution	No CAS	Unique timestamp	ABA-safe by default
Memory reclamation	Custom scheme	Custom scheme	Custom scheme
Progress guarantee of dequeue	Wait-free	Wait-free	Wait-free
Theoretical performance model of dequeue	2A	$\Theta(\log n)R$ + $\Theta(\log n)A$	$\Theta(1)R$ + $\Theta(n)A$
Progress guarantee of enqueue	Wait-free	Wait-free	Wait-free
Theoretical performance model of enqueue	2R	$\Theta(\log n)R$ + $\Theta(\log n)A$	$\Theta(1)R$
Number of elements	Bounded	Bounded (2)	Bounded (2)

Table 3: Characteristic summary of existing distributed FIFO queues.

R stands for remote operations and **A** stands for atomic operations

- (1) the **baseline SPSC** refers to the SPSC we introduce in Section 4.2, the reason we have to qualify **dLTQueue** and **Slotqueue** with a specific SPSC implementation because **dLTQueue** and **Slotqueue** are in fact “SPSC wrappers” that turn any SPSCs to MPSCs.
- (2) The “bounded” property is not inherent for **dLTQueue** and **Slotqueue** “wrappers”, they are bounded because the **baseline SPSC** is bounded.

3.2.1 Fastqueue - BCL’s MP/MC FIFO data structure

3.2.2 dLTQueue & Slotqueue

Chapter IV Distributed MPSC queues

Based on the MPSC queue algorithms we have surveyed in Chapter III, we propose two wait-free distributed MPSC queue algorithms:

- dLTQueue (Section 4.3) is a direct modification of the original LTQueue [7] without any usage of LL/SC, adapted for distributed environment.
- Slotqueue (Section 4.4) is inspired by the timestamp-refreshing idea of LTQueue [7] and repeated-rescan of Jiffy [8]. Although it still bears some resemblance to LTQueue, we believe it to be more optimized for distributed context.

MPSC queues	dLTQueue	Slotqueue
Correctness	Linearizable	Linearizable
Progress guarantee of dequeue	Wait-free	Wait-free
Progress guarantee of enqueue	Wait-free	Wait-free
ABA solution	Unique timestamp	ABA-safe by default
Memory reclamation	Custom scheme	Custom scheme

Table 4: Characteristic summary of our proposed distributed MPSC queues.
 n is the number of enqueueers, R stands for **remote operation** and A stands for **atomic operation**.

In this section, we present our proposed distributed MPSC queues in detail. Any other discussions about theoretical aspects of these algorithms such as linearizability, progress guarantee, time complexity are deferred to Chapter V.

In our description, we assume that each process in our program is assigned a unique number as an identifier, which is termed as its **rank**. The numbers are taken from the range of $[0, \text{size} - 1]$, with size being the number of processes in our program.

4.1 Distributed one-sided-communication primitives in our distributed algorithm specification

Although we use MPI-3 RMA to implement these algorithms, the algorithm specifications themselves are not inherently tied to MPI-3 RMA interfaces. For clarity and convenience in specification, we define the following distributed primitives used in our pseudocode.

remote<T>: A distributed shared variable of type T. The process that physically stores the variable in its local memory is referred to as the **host**. This represents data that can be accessed or modified remotely by other processes.

void aread_sync(remote<T> src, T* dest): Issue a synchronous read of the distributed variable src and stores its value into the local memory location pointed to by dest. The read is guaranteed to be completed when the function returns.

`void aread_sync(remote<T*> src, int index, T* dest):` Issue a synchronous read of the element at position `index` within the distributed array `src` (where `src` is a pointer to a remotely hosted array of type `T`) and stores the value into the local memory location pointed to by `dest`. The read is guaranteed to be completed when the function returns.

`void awrite_sync(remote<T> dest, T* src):` Issue a synchronous write of the value at the local memory location pointed to by `src` into the distributed variable `dest`. The write is guaranteed to be completed when the function returns.

`void awrite_sync(remote<T*> dest, int index, T* src):` Issue a synchronous write of the value at the local memory location pointed to by `src` into the element at position `index` within the distributed array `dest` (where `dest` is a pointer to a remotely hosted array of type `T`). The write is guaranteed to be completed when the function returns.

`void aread_async(remote<T> src, T* dest):` Issue an asynchronous read of the distributed variable `src` and initiate the transfer of its value into the local memory location pointed to by `dest`. The operation may not be completed when the function returns.

`void aread_async(remote<T*> src, int index, T* dest):` Issue an asynchronous read of the element at position `index` within the distributed array `src` (where `src` is a pointer to a remotely hosted array of type `T`) and initiate the transfer of its value into the local memory location pointed to by `dest`. The operation may not be completed when the function returns.

`void awrite_async(remote<T> dest, T* src):` Issue an asynchronous write of the value at the local memory location pointed to by `src` into the distributed variable `dest`. The operation may not be completed when the function returns.

`void awrite_async(remote<T*> dest, int index, T* src):` Issue an asynchronous write of the value at the local memory location pointed to by `src` into the element at position `index` within the distributed array `dest` (where `dest` is a pointer to a remotely hosted array of type `T`). The operation may not be completed when the function returns.

`void flush(remote<T> src):` Ensure that all read and write operations on the distributed variable `src` (or its associated array) issued before this function call are fully completed by the time the function returns.

`bool compare_and_swap_sync(remote<T> dest, T old_value, T new_value):` Issue a synchronous compare-and-swap operation on the distributed variable `dest`. The operation atomically compares the current value of `dest` with `old_value`. If they are equal, the value of `dest` is replaced with `new_value`; otherwise, no change is made. The operation is guaranteed to be completed when the function returns, ensuring that the update (if any) is visible to all processes. The type `T` must be a data type with a size of 1, 2, 4, or 8 bytes.

`bool compare_and_swap_sync(remote<T*> dest, int index, T old_value, T new_value):` Issue a synchronous compare-and-swap operation on the element at posi-

tion index within the distributed array `dest` (where `dest` is a pointer to a remotely hosted array of type τ). The operation atomically compares the current value of the element at `dest[index]` with `old_value`. If they are equal, the element at `dest[index]` is replaced with `new_value`; otherwise, no change is made. The operation is guaranteed to be completed when the function returns, ensuring that the update (if any) is visible to all processes. The type τ must be a data type with a size of 1, 2, 4, or 8.

τ `fetch_and_add_sync(remote< τ > dest, τ inc)`: Issue a synchronous fetch-and-add operation on the distributed variable `dest`. The operation atomically adds the value `inc` to the current value of `dest`, returning the original value of `dest` (before the addition) to the calling process. The update to `dest` is guaranteed to be completed and visible to all processes when the function returns. The type τ must be an integral type with a size of 1, 2, 4, or 8 bytes.

4.2 A simple baseline distributed SPSC

The two algorithms we propose here both utilize a distributed SPSC data structure, which we will present first. For implementation simplicity, we present a bounded SPSC, effectively make our proposed algorithms support only a bounded number of elements. However, one can trivially substitute another distributed unbounded SPSC to make our proposed algorithms support an unbounded number of elements, as long as this SPSC supports the same interface as ours.

Placement-wise, all shared data in this SPSC is hosted on the enqueuer.

Types

| `data_t` = The type of data stored

Shared variables

| `First: remote<uint64_t>`
| The index of the last undequeued entry. Hosted at the enqueuer.
| `Last: remote<uint64_t>`
| The index of the last unenqueued entry. Hosted at the enqueuer.
| `Data: remote<data_t*>`
| An array of `data_t` of some known capacity. Hosted at the enqueuer.

Enqueuer-local variables

| `Capacity`: A read-only value indicating the capacity of the SPSC
| `First_buf`: The cached value of `First`
| `Last_buf`: The cached value of `Last`

Dequeuer-local variables

| `Capacity`: A read-only value indicating the capacity of the SPSC
| `First_buf`: The cached value of `First`
| `Last_buf`: The cached value of `Last`

Enqueuer initialization

| Initialize `First` and `Last` to 0
| Initialize `Capacity`

| Allocate array in `Data`
| `First_buf = Last_buf = 0`

Dequeuer initialization

Initialize Capacity

```
First_buf = Last_buf = 0
```

The procedures of the enqueueer are given as follows.

Procedure 4: bool spsc_enqueue(data_t v)

```

1 new_last = Last_buf + 1
2 if (new_last - First_buf > Capacity)
3     | aread_sync(First, &First_buf)
4     | if (new_last - First_buf > Capacity)
5     |     | return false
6 arewrite_sync(Data, Last_buf % Capacity, &v)
7 arewrite_sync(Last, &new_last)
8 Last_buf = new_last
9 return true

```

spsc_enqueue first computes the new Last value (line 1). If the queue is full as indicated by the difference between the new Last value and First-buf (line 2), there can still be the possibility that some elements have been dequeued but First-buf hasn't been synced with First yet, therefore, we first refresh the value of First-buf by fetching from First (line 3). If the queue is still full (line 4), we signal failure (line 5). Otherwise, we proceed to write the enqueued value to the entry at Last_buf % Capacity (line 6), increment Last (line 7), update the value of Last_buf (line 8) and signal success (line 9).

Procedure 5: `bool spsc_readFronte(data_t* output)`

```

10 if (First_buf >= Last_buf)
11 | return false
12 aread_sync(First, &First_buf)
13 if (First_buf >= Last_buf)
14 | return false
15 aread_sync(Data, First_buf % Capacity, output)
16 return true

```

`spsc_readFront`, first checks if the SPSC is empty based on the difference between `First_buf` and `Last_buf` (line 10). Note that if this check fails, we signal failure immediately (line 11) without refetching either `First` or `Last`. This suffices because `Last` cannot be out-of-sync with `Last_buf` as we're the enqueuer and `First` can only increase since the last refresh of `First_buf`, therefore, if we refresh `First` and `Last`, the condition on line 10 would return `false` anyways. If the SPSC is not empty, we refresh

First and re-perform the empty check (line 12-14). If the SPSC is again not empty, we read the queue entry at `First_buf % Capacity` into output (line 15) and signal success (line 16).

The procedures of the dequeuer are given as follows.

Procedure 6: `bool spsc_dequeue(data_t* output)`

```
15 new_first = First_buf + 1
16 if (new_first > Last_buf)
17     aread_sync(Last, &Last_buf)
18     if (new_first > Last_buf)
19         | return false
20 aread_sync(Data, First_buf % Capacity, output)
21 awrite_sync(First, &new_first)
22 First_buf = new_first
23 return true
```

`spsc_dequeue` first computes the new `First` value (line 15). If the queue is empty as indicated by the difference the new `First` value and `Last_buf` (line 16), there can still be the possibility that some elements have been enqueued but `Last_buf` hasn't been synced with `Last` yet, therefore, we first refresh the value of `Last_buf` by fetching from `Last` (line 17). If the queue is still empty (line 18), we signal failure (line 19). Otherwise, we proceed to read the top value at `First_buf % Capacity` (line 20) into output, increment `First` (line 21) - effectively dequeue the element, update the value of `First_buf` (line 22) and signal success (line 23).

Procedure 7: `bool spsc_readFrontd(data_t* output)`

```
24 if (First_buf >= Last_buf)
25     aread_sync(Last, &Last_buf)
26     if (First_buf >= Last_buf)
27         | return false
28 aread_sync(Data, First_buf % Capacity, output)
29 return true
```

`spsc_readFrontd` first checks if the SPSC is empty based on the difference between `First_buf` and `Last_buf` (line 24). If this check fails, we refresh `Last_buf` (line 25) and recheck (line 26). If the recheck fails, signal failure (line 27). If the SPSC is not empty, we read the queue entry at `First_buf % Capacity` into output (line 28) and signal success (line 29).

4.3 dLTQueue - Modified distributed LTQueue without LL/SC

This algorithm presents our most straightforward effort to port LTQueue [7] to distributed context. The main challenge is that LTQueue uses LL/SC as the universal atomic instruction and also an ABA solution, but LL/SC is not available in distributed programming environments. We have to replace any usage of LL/SC in the original LTQueue algorithm. Compare-and-swap is unavoidable in distributed MPSC queues, so we use the well-known monotonic timestamp scheme to guard against ABA problem.

4.3.1 Data structure & Algorithm overview

4.3.2 Data structure

The structure of our dLTQueue is shown as in Image 2.

We differentiate between 2 types of nodes: **enqueueer nodes** (represented as the rectangular boxes at the bottom of Image 2) and normal **tree nodes** (represented as the circular boxes in Image 2).

Each enqueueer node corresponds to an enqueueer. Each time the local SPSC is enqueued with a value, the enqueueer timestamps the value using a distributed counter shared by all enqueueers. An enqueueer node stores the SPSC local to the corresponding enqueueer and a `min_timestamp` value which is the minimum timestamp inside the local SPSC.

Each tree node stores the rank of an enqueueer process. This rank corresponds to the enqueueer node with the minimum timestamp among the node's children's ranks. The tree node that's attached to an enqueueer node is called a **leaf node**, otherwise, it's called an **internal node**.

Note that if a local SPSC is empty, the `min_timestamp` variable of the corresponding enqueueer node is set to `MAX_TIMESTAMP` and the corresponding leaf node's rank is set to `DUMMY_RANK`.

Placement-wise:

- The **enqueueer nodes** are hosted at the corresponding **enqueueer**.
- All the **tree nodes** are hosted at the **dequeueer**.
- The distributed counter, which the enqueueers use to timestamp their enqueued value, is hosted at the **dequeueer**.

4.3.3 Algorithm

Below is the types utilized in dLTQueue.

Types

`data_t` = The type of the data to be stored

`spsc_t` = The type of the SPSC, this is assumed to be the distributed SPSC in Section 4.2

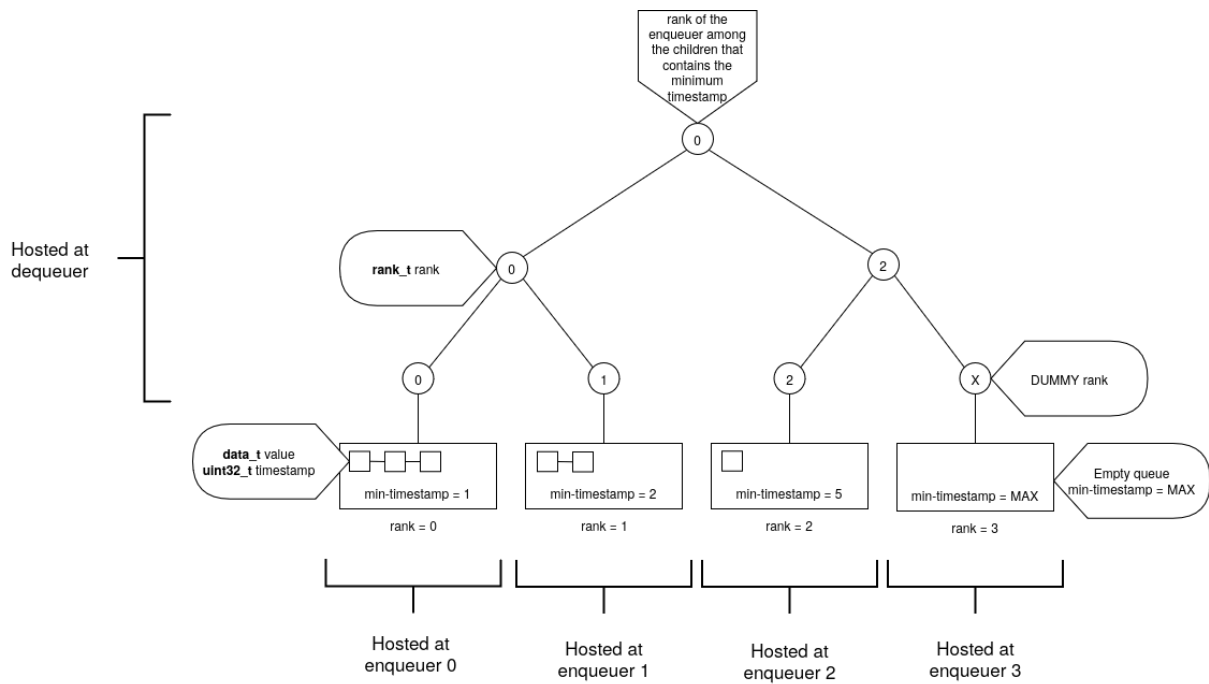


Image 2: dLTQueue's structure.

rank_t = The type of the rank of an enqueueer process tagged with a unique timestamp (version) to avoid ABA problem

```
struct
| value: uint32_t
| version: uint32_t
end
```

timestamp_t = The type of the timestamp tagged with a unique timestamp (version) to avoid ABA problem

```
struct
| value: uint32_t
| version: uint32_t
end
```

node_t = The type of a tree node

```
struct
| rank: rank_t
end
```

The shared variables in our LTQueue version are as followed.

Note that we have described a very specific and simple way to organize the tree nodes in dLTQueue in a min-heap-like array structure hosted on the sole dequeuer. We will resume our description of the related tree-structure procedures parent() (Procedure 9), children() (Procedure 10), leafNodeIndex() (Procedure 11) with this representation in mind. However, our algorithm doesn't strictly require this representation and can be

substituted with other more-optimized representations & distributed placements, as long as the similar tree-structure procedures are supported.

Shared variables

Counter: remote<uint64_t>

| A distributed counter shared by the enqueueers. Hosted at the dequeuer.

Tree_size: uint64_t

| A read-only variable storing the number of tree nodes present in the dLTQueue.

Nodes: remote<node_t>

| An array with Tree_size entries storing all the tree nodes present in the dLTQueue shared by all processes.

| Hosted at the dequeuer.

| This array is organized in a similar manner as a min-heap: At index 0 is the root node. For every index $i > 0$, $\lfloor \frac{i-1}{2} \rfloor$ is the index of the parent of node i . For every index $i > 0$, $2i + 1$ and $2i + 2$ are the indices of the children of node i .

Dequeuer_rank: uint32_t

| The rank of the dequeuer process. This is read-only.

Timestamps: A read-only **array** [0..size - 2] of remote<timestamp_t>, with size being the number of processes.

| The entry at index i corresponds to the Min_timestamp distributed variable at the enqueueer with an order of i .

Similar to the fact that each process in our program is assigned a rank, each enqueueer process in our program is assigned an **order**. The following procedure computes an enqueueer's order based on its rank:

Procedure 8: uint32_t enqueueerOrder(uint32_t enqueueer_rank)

```
1 return enqueueer_rank > Dequeuer_rank ? enqueueer_rank - 1 : enqueueer_rank
```

This procedure is rather straightforward: Each enqueueer is assigned an order in the range [0, size - 2], with size being the number of processes and the total ordering among the enqueueers based on their ranks is the same as the total ordering among the enqueueers based on their orders.

Enqueuer-local variables

Enqueuer_count: uint64_t
| The number of enqueueers.
Self_rank: uint32_t
| The rank of the current enqueueer process.
Min_timestamp:
remote<timestamp_t>
SpSC: spsc_t
| This SPSC is synchronized with the dequeuer.

Dequeuer-local variables

Enqueuer_count: uint64_t
| The number of enqueueers.
SpSCs: **array** of spsc_t with Enqueuer_count entries.
| The entry at index i corresponds to the SpSC at the enqueueer with an order of i .

Enqueuer initialization

Initialize Enqueuer_count, Self_rank and Dequeuer_rank.
Initialize SpSC to the initial state.
Initialize Min_timestamp to timestamp_t {MAX_TIMESTAMP, 0}.

Dequeuer initialization

Initialize Enqueuer_count, Self_rank and Dequeuer_rank.
Initialize Counter to 0.
Initialize Tree_size to Enqueuer_count * 2.
Initialize Nodes to an array with Tree_size entries. Each entry is initialized to node_t {DUMMY_RANK}.
Initialize SpSCs, synchronizing each entry with the corresponding enqueueer.
Initialize Timestamps, synchronizing each entry with the corresponding enqueueer.

We first present the tree-structure utility procedures that are shared by both the enqueueer and the dequeuer:

Procedure 9: uint32_t parent(uint32_t index)

2 **return** (index - 1) / 2

parent returns the index of the parent tree node given the node with index index. These indices are based on the shared Nodes array. Based on how we organize the Nodes array, the index of the parent tree node of index is $(\text{index} - 1) / 2$.

Procedure 10: `vector<uint32_t> children(uint32_t index)`

```
3 left_child = index * 2 + 1
4 right_child = left_child + 1
5 res = vector<uint32_t>()
6 if (left_child >= Tree_size)
7 | return res
8 res.push(left_child)
9 if (right_child >= Tree_size)
10 | return res
11 res.push(right_child)
12 return res
```

Similarly, `children` returns all indices of the child tree nodes given the node with index `index`. These indices are based on the shared `Nodes` array. Based on how we organize the `Nodes` array, these indices can be either `index * 2 + 1` or `index * 2 + 2`.

Procedure 11: `uint32_t leafNodeIndex(uint32_t enqueue_rank)`

```
13 return Tree_size + enqueueOrder(enqueue_rank)
```

`leafNodeIndex` returns the index of the leaf node that's logically attached to the enqueue node with rank `enqueue_rank` as in Image 2.

The followings are the enqueue procedures.

Procedure 12: `bool enqueue(data_t value)`

```
14 timestamp = fetch_and_add_sync(Counter, 1)
15 spsc_enqueue(&Spsc, (value, timestamp))
16 propagate_e()
```

To enqueue a value, `enqueue` first obtains a count by FAA the distributed counter `Counter` (line 14). Then, we enqueue the data tagged with the timestamp into the local SPSC (line 15). Finally, `enqueue` propagates the changes by invoking `propagatee()` (line 16).

Procedure 13: void propagate_e()

```
18 if (!refreshTimestampe())
19 | refreshTimestampe()
20 if (!refreshLeafe())
21 | refreshLeafe()
22 current_node_index = leafNodeIndex(Self_rank)
23 repeat
24 | current_node_index = parent(current_node_index)
25 | if (!refreshe(current_node_index))
26 | | refreshe(current_node_index)
27 until current_node_index == 0
```

The propagate_e procedure is responsible for propagating SPSC updates up to the root node as a way to notify other processes of the newly enqueued item. It is split into 3 phases: Refreshing of Min_timestamp in the enqueuer node (line 18-19), refreshing of the enqueuer's leaf node (line 20-21), refreshing of internal nodes (line 23-27). On line 20-27, we refresh every tree node that lies between the enqueuer node and the root node.

Procedure 14: bool refreshTimestamp_e()

```
28 min_timestamp = timestamp_t {}
29 aread_sync(Min_timestamp, &min_timestamp)
30 {old-timestamp, old-version} = min_timestamp
31 front = (data_t {}, timestamp_t {})
32 is_empty = spsc_readFront(Spsc, &front)
33 if (is_empty)
34 | return compare_and_swap_sync(Min_timestamp,
35 | timestamp_t {old-timestamp, old-version},
36 | timestamp_t {MAX_TIMESTAMP, old-version + 1})
35 else
36 | return compare_and_swap_sync(Min_timestamp,
37 | timestamp_t {old-timestamp, old-version},
38 | timestamp_t {front.timestamp, old-version + 1})
```

The refreshTimestamp_e procedure is responsible for updating the Min_timestamp of the enqueuer node. It simply looks at the front of the local SPSC (line 310 and CAS Min_timestamp accordingly (line 33-36).

Procedure 15: `bool refreshNodee(uint32_t current_node_index)`

```
37 current_node = node_t {}
38 aread_sync(Nodes, current_node_index, &current_node)
39 {old_rank, old_version} = current_node.rank
40 min_rank = DUMMY_RANK
41 min_timestamp = MAX_TIMESTAMP
42 for child_node_index in children(current_node)
43     child_node = node_t {}
44     aread_sync(Nodes, child_node_index, &child_node)
45     {child_rank, child_version} = child_node
46     if (child_rank == DUMMY_RANK) continue
47     child_timestamp = timestamp_t {}
48     aread_sync(Timestamps[enqueuerOrder(child_rank)], &child_timestamp)
49     if (child_timestamp < min_timestamp)
50         min_timestamp = child_timestamp
51         min_rank = child_rank
    return compare_and_swap_sync(Nodes, current_node_index,
52 node_t {rank_t {old_rank, old_version}},
    node_t {rank_t {min_rank, old_version + 1}})
```

The `refreshNodee` procedure is responsible for updating the ranks of the internal nodes affected by the enqueue. It loops over the children of the current internal nodes (line 42). For each child node, we read the rank stored in it (line 44), if the rank is not `DUMMY_RANK`, we proceed to read the value of `Min_timestamp` of the enqueuer node with the corresponding rank (line 48). At the end of the loop, we obtain the rank stored inside one of the child nodes that has the minimum timestamp stored in its enqueuer node (line 50-51). We then try to CAS the rank inside the current internal node to this rank.

Procedure 16: `bool refreshLeafe()`

```
53 leaf_node_index = leafNodeIndex(Self_rank)
54 leaf_node = node_t {}
55 aread_sync(Nodes, leaf_node_index, &leaf_node)
56 {old_rank, old_version} = leaf_node.rank
57 min_timestamp = timestamp_t {}
58 aread_sync(Min_timestamp, &min_timestamp)
59 timestamp = min_timestamp.timestamp
    return compare_and_swap_sync(Nodes, leaf_node_index,
60 node_t {rank_t {old_rank, old_version}},
    node_t {timestamp == MAX ? DUMMY_RANK : Self_rank, old_version + 1})
```

The `refreshLeafe` procedure is responsible for updating the rank of the leaf node affected by the enqueue. It simply reads the value of `Min_timestamp` of the enqueuer node it's logically attached to (line 58) and CAS the leaf node's rank accordingly (line 60).

The followings are the dequeuer procedures.

Procedure 17: `bool dequeue(data_t* output)`

```
61 root_node = node_t {}
62 aread_sync(Nodes, 0, &root_node)
63 {rank, version} = root_node.rank
64 if (rank == DUMMY_RANK) return false
65 output_with_timestamp = (data_t {}, timestamp_t {})
66 if (!spsc_dequeue(&Spscs[enqueuerOrder(rank)]),
    &output_with_timestamp))
67 | return false
68 *output = output_with_timestamp.data
69 propagated(rank)
70 return true
```

To dequeue a value, `dequeue` reads the rank stored inside the root node (line 62). If the rank is `DUMMY_RANK`, the MPSC queue is treated as empty and failure is signaled (line 64). Otherwise, we invoke `spsc_dequeue` on the SPSC of the enqueuer with the obtained rank (line 66). We then extract out the real data and set it to `output` (line 68). We finally propagate the dequeue from the enqueuer node that corresponds to the obtained rank (line 69) and signal success (line 70).

Procedure 18: `void propagated(uint32_t enqueuer_rank)`

```
71 if (!refreshTimestampd(enqueuer_rank))
72 | refreshTimestampd(enqueuer_rank)
73 if (!refreshLeafd(enqueuer_rank))
74 | refreshLeafd(enqueuer_rank)
75 current_node_index = leafNodeIndex(enqueuer_rank)
76 repeat
77 | current_node_index = parent(current_node_index)
78 | if (!refreshd(current_node_index))
79 | | refreshd(current_node_index)
80 until current_node_index == 0
```

The `propagated` procedure is similar to `propagatee`, with appropriate changes to accommodate the dequeuer.

Procedure 19: bool refreshTimestamp_d(uint32_t enqueue_rank)

```
81 enqueue_order = enqueueOrder(enqueue_rank)
82 min_timestamp = timestamp_t {}
83 aread_sync(Timestamps, enqueue_order, &min_timestamp)
84 {old-timestamp, old-version} = min_timestamp
85 front = (data_t {}, timestamp_t {})
86 is_empty = spsc_readFront(&Spscs[enqueue_order], &front)
87 if (is_empty)
    | return compare_and_swap_sync(Timestamps, enqueue_order,
88 | timestamp_t {old-timestamp, old-version},
    | timestamp_t {MAX_TIMESTAMP, old-version + 1})
89 else
    | return compare_and_swap_sync(Timestamps, enqueue_order,
90 | timestamp_t {old-timestamp, old-version},
    | timestamp_t {front.timestamp, old-version + 1})
```

The refreshTimestamp_d procedure is similar to refreshTimestamp_e, with appropriate changes to accommodate the dequeuer.

Procedure 20: bool refreshNode_d(uint32_t current_node_index)

```
91 current_node = node_t {}
92 aread_sync(Nodes, current_node_index, &current_node)
93 {old-rank, old-version} = current_node.rank
94 min_rank = DUMMY_RANK
95 min_timestamp = MAX_TIMESTAMP
96 for child_node_index in children(current_node)
97 | child_node = node_t {}
98 | aread_sync(Nodes, child_node_index, &child_node)
99 | {child_rank, child_version} = child_node
100 | if (child_rank == DUMMY_RANK) continue
101 | child_timestamp = timestamp_t {}
102 | aread_sync(Timestamps[enqueueOrder(child_rank)], &child_timestamp)
103 | if (child_timestamp < min_timestamp)
104 | | min_timestamp = child_timestamp
105 | | min_rank = child_rank
    | return compare_and_swap_sync(Nodes, current_node_index,
106 node_t {rank_t {old_rank, old_version}},
    node_t {rank_t {min_rank, old_version + 1}})
```

The refreshNode_d procedure is similar to refreshNode_e , with appropriate changes to accommodate the dequeuer.

Procedure 21: $\text{bool refreshLeaf}_d(\text{uint32_t enqueue_rank})$

```
107 leaf_node_index = leafNodeIndex(enqueue_rank)
108 leaf_node = node_t {}
109 aread_sync(Nodes, leaf_node_index, &leaf_node)
110 {old_rank, old_version} = leaf_node.rank
111 min_timestamp = timestamp_t {}
112 aread_sync(Timestamps, enqueueOrder(enqueue_rank), &min_timestamp)
113 timestamp = min_timestamp.timestamp
    return compare_and_swap_sync(Nodes, leaf_node_index,
114 node_t {rank_t {old_rank, old_version}},
    node_t {timestamp == MAX ? DUMMY_RANK : Self_rank, old_version + 1})
```

The refreshLeaf_d procedure is similar to refreshLeaf_e , with appropriate changes to accommodate the dequeuer.

4.4 Slotqueue - Optimized dLTQueue for distributed context

4.4.1 Motivation

Even though the straightforward dLTQueue algorithm we have ported in Section 4.3 pretty much preserve the original algorithm's characteristics, i.e. wait-freedom and time complexity of $\Theta(\log n)$ for both enqueue and dequeue operations (which we will prove in Chapter V), we have to be aware that this is $\Theta(\log n)$ remote operations, which is potentially expensive and a bottleneck in the algorithm.

Therefore, to be more suitable for distributed context, we propose a new algorithm that's inspired by LTQueue, in which both enqueue and dequeue only perform a constant number of remote operations, at the cost of dequeue having to perform $\Theta(n)$ local operations, where n is the number of enqueueers. Because remote operations are much more expensive, this might be a worthy tradeoff.

4.4.2 Data structure & Algorithm overview

4.4.3 Data structure

The structure of Slotqueue is shown as in Figure 18.

Each enqueueer hosts a distributed SPSC as in dLTQueue (Section 4.3). The enqueueer when enqueues a value to its local SPSC will timestamp the value using a distributed counter hosted at the dequeuer.

Additionally, the dequeuer hosts an array whose entries each corresponds with an enqueueer. Each entry stores the minimum timestamp of the local SPSC of the corresponding enqueueer.

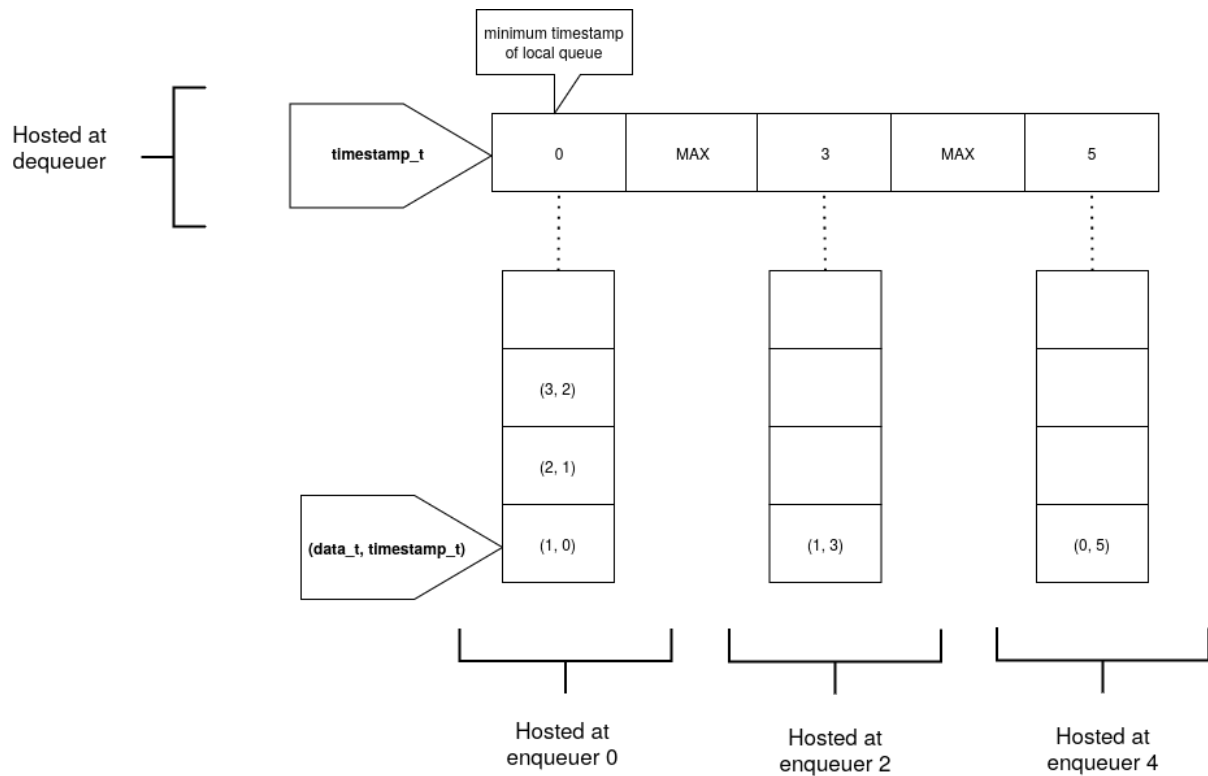


Figure 18: Basic structure of Slotqueue.

4.4.4 Algorithm

We first introduce the types and shared variables utilized in Slotqueue.

Types

data_t = The type of data stored

timestamp_t = uint64_t

spsc_t = The type of the SPSC each enqueueer uses, this is assumed to be the distributed SPSC in Section 4.2

Shared variables

Slots: remote<timestamp_t*>

An array of timestamp_t with the number of entries equal to the number of enqueueers.

Hosted at the dequeuer.

Counter: remote<uint64_t>

A distributed counter.

Hosted at the dequeuer.

Dequeuer_rank: uint32_t

| | The rank of the dequeuer process. This is read-only.

Similar to the idea of assigning an order to each enqueueer in dLTQueue, the following procedure computes an enqueueer's order based on its rank:

Procedure 22: uint64_t enqueueerOrder(uint64_t enqueueer_rank)

```
1 return enqueueer_rank > Dequeueer_rank ? enqueueer_rank - 1 : enqueueer_rank
```

Again, each enqueueer is assigned an order in the range $[0, \text{size} - 2]$, with size being the number of processes and the total ordering among the enqueueers based on their ranks is the same as the total ordering among the enqueueers based on their orders.

Reversely, enqueueerRank computes an enqueueer's rank given its order.

Procedure 23: uint64_t enqueueerRank(uint64_t enqueueer_order)

```
2 return enqueueer_order >= Dequeueer_rank ? enqueueer_order + 1 : enqueueer_order
```

Enqueueer-local variables

Dequeueer_rank: uint64_t	
Enqueueer_count: uint64_t	The number of enqueueers.
Self_rank: uint32_t	
	The rank of the current enqueueer process.
Spssc: spssc_t	
	This SPSC is synchronized with the dequeuer.

Dequeuer-local variables

Dequeueer_rank: uint64_t	
Enqueueer_count: uint64_t	The number of enqueueers.
Spscs: array of spssc_t with Enqueueer_count entries.	
	The entry at index i corresponds to the Spssc at the enqueueer with an order of i .

The enqueueer operations are given as follows.

Procedure 24: bool enqueue(data_t v)

```
3 timestamp = fetch_and_add_sync(Counter)
4 if (!spssc_enqueue(&Spssc, (v, timestamp))) return false
5 if (!refreshEnqueue(timestamp))
6 | refreshEnqueue(timestamp)
7 return true
```

To enqueue a value, enqueue first obtains a timestamp by FAA-ing the distributed counter (line 3). It then tries to enqueue the value tagged with the timestamp (line 4). At line 5-6, the enqueuer tries to refresh its slot's timestamp.

Procedure 25: `bool refreshEnqueue(timestamp_t ts)`

```
8 enqueue_order = enqueueOrder(Self_rank)
9 front = (data_t {}, timestamp_t {})
10 success = spsc_readFront(Spsc, &front)
11 new_timestamp = success ? front.timestamp : MAX_TIMESTAMP
12 if (new_timestamp != ts)
13 | return true
14 old_timestamp = timestamp_t {}
15 aread_sync(&Slots, enqueue_order, &old_timestamp)
16 success = spsc_readFront(Spsc, &front)
17 new_timestamp = success ? front.timestamp : MAX_TIMESTAMP
18 if (new_timestamp != ts)
19 | return true
    return compare_and_swap_sync(Slots, enqueue_order,
20     old_timestamp,
        new_timestamp)
```

`refreshEnqueue`'s responsibility is to refresh the timestamp stores in the enqueuer's slot to potentially notify the dequeuer of its newly-enqueued element. It first reads the current front element (line 10). If the SPSC is empty, the new timestamp is set to `MAX_TIMESTAMP`, otherwise, the front element's timestamp (line 11). If it finds that the front element's timestamp is different from the timestamp `ts` it returns `true` immediately (line 9-13). Otherwise, it reads its slot's old timestamp (line 14) and re-reads the current front element in the SPSC (line 16) to update the new timestamp. Note that similar to line 13, `refreshEnqueue` immediately succeeds if the new timestamp is different from the timestamp `ts` of the element it enqueues (line 19). Otherwise, it tries to CAS its slot's timestamp with the new timestamp (line 20).

The dequeuer operations are given as follows.

Procedure 26: bool dequeue(data_t* output)

```
21 rank = readMinimumRank()
22 if (rank == DUMMY_RANK)
23 | return false
24 output_with_timestamp = (data_t {}, timestamp_t {})
25 if (!spsc_dequeue(Spsc, &output_with_timestamp))
26 | return false
27 *output = output_with_timestamp.data
28 if (!refreshDequeue(rank))
29 | refreshDequeue(rank)
30 return true
```

To dequeue a value, dequeue first reads the rank of the enqueuer whose slot currently stores the minimum timestamp (line 21). If the obtained rank is DUMMY_RANK, failure is signaled (line 22-23). Otherwise, it tries to dequeue the SPSC of the corresponding enqueuer (line 25). It then tries to refresh the enqueuer's slot's timestamp to potentially notify the enqueuer of the dequeue (line 28-29). It then signals success (line 30).

Procedure 27: uint64_t readMinimumRank()

```
31 buffered_slots = timestamp_t[Enqueuer_count] {}
32 for index in 0..Enqueuer_count
33 | aread_async(Slots, index, &buffered_slots[index])
34 flush(Slots)
35 if every entry in buffered_slots is MAX_TIMESTAMP
36 | return DUMMY_RANK
37 for index in 0..Enqueuer_count
38 | aread_async(Slots, index, &buffered_slots[index])
39 flush(Slots)
40 rank = DUMMY_RANK
41 min_timestamp = MAX_TIMESTAMP
42 for index in 0..Enqueuer_count
43 | timestamp = buffered_slots[index]
44 | if (min_timestamp < timestamp)
45 | | rank = enqueueRank(index)
46 | | min_timestamp = timestamp
47 return rank
```

readMinimumRank's main responsibility is to return the rank of the enqueuer from which we can safely dequeue next. It first creates a local buffer to store the value

read from slots (line 31). It then performs 2 scans of slots and read every entry into buffered_slots (line 32-37). If the first scan finds only MAX_TIMESTAMPS, DUMMY_RANK is returned (line 36). From there, based on buffered_slots, it returns the rank of the enqueueer whose buffered slot stores the minimum timestamp (line 42-47).

Procedure 28: refreshDequeue(rank: int) **returns** bool

```
48 enqueueer_order = enqueueerOrder(rank)
49 old_timestamp = timestamp_t {}
50 aread_sync(&Slots, enqueueer_order, &old_timestamp)
51 front = (data_t {}, timestamp_t {})
52 success = spsc_readFront(SpSCs[enqueueer_order], &front)
53 new_timestamp = success ? front.timestamp : MAX_TIMESTAMP
   return compare_and_swap_sync(Slots, enqueueer_order,
54     old_timestamp,
       new_timestamp)
```

refreshDequeue's responsibility is to refresh the timestamp of the just-dequeued enqueueer to notify the enqueueer of the dequeue. It first reads the old timestamp of the slot (line 50) and the front element (line 52). If the SPSC is empty, the new timestamp is set to MAX_TIMESTAMP, otherwise, it's the front element's timestamp (line 53). It finally tries to CAS the slot with the new timestamp (line 54).

Chapter V Theoretical aspects

This section discusses the correctness and progress guarantee properties of the distributed MPSC queue algorithms introduced in Chapter IV. We also provide a theoretical performance model of these algorithms to predict how well they scale to multiple nodes.

5.1 Terminology

In this section, we introduce some terminology that we will use throughout our proofs.

Definition 5.1.1 In an SPSC/MPSC queue, an enqueue operation e is said to **match** a dequeue operation d if d returns the value that e enqueues. Similarly, d is said to **match** e . In this case, both e and d are said to be **matched**.

Definition 5.1.2 In an SPSC/MPSC queue, an enqueue operation e is said to be **unmatched** if no dequeue operation **matches** it.

Definition 5.1.3 In an SPSC/MPSC queue, a dequeue operation d is said to be **unmatched** if no enqueue operation **matches** it, in other word, d returns false.

5.2 Preliminaries

In this section, we formalize the notion of correct concurrent algorithms and harmless ABA problem. We will base our proofs on these formalisms to prove their correctness. We also provide a simple way to theoretically model our queues' performance.

5.2.1 Linearizability

Linearizability is a criteria for evaluating a concurrent algorithm's correctness. This is the model we use to prove our algorithm's correctness. Our formalization of linearizability is equivalent to that of [9] by Herlihy and Shavit. However, there are some differences in our terminology.

For a concurrent object S , we can call some methods on S concurrently. A method call on the object S is said to have an **invocation event** when it starts and a **response event** when it ends.

Definition 5.2.1.1 An **invocation event** is a triple $(S, t, args)$, where S is the object the method is invoked on, t is the timestamp of when the event happens and $args$ is the arguments passed to the method call.

Definition 5.2.1.2 A **response event** is a triple (S, t, res) , where S is the object the method is invoked on, t is the timestamp of when the event happens and res is the results of the method call.

Definition 5.2.1.3 A **method call** is a tuple of (i, r) where i is an invocation event and r is a response event or the special value \perp indicating that its response event hasn't happened yet. A well-formed **method call** should have a response event with a larger timestamp than its invocation event or the response event hasn't happened yet.

Definition 5.2.1.4 A **method call** is **pending** if its invocation event is \perp .

Definition 5.2.1.5 A **history** is a set of well-formed **method calls**.

Definition 5.2.1.6 An extension of **history** H is a **history** H' such that any pending method call is given a response event.

We can define a **strict partial order** on the set of well-formed method calls:

Definition 5.2.1.7 \rightarrow is a relation on the set of well-formed method calls. With two method calls X and Y , we have $X \rightarrow Y \Leftrightarrow X$'s response event is not \perp and its response timestamp is not greater than Y 's invocation timestamp.

Definition 5.2.1.8 Given a **history** H , \rightarrow_H is a relation on H such that for two method calls X and Y in H , $X \rightarrow_H Y \Leftrightarrow X \rightarrow Y$.

Definition 5.2.1.9 A **sequential history** H is a **history** such that \rightarrow_H is a total order on H .

Now that we have formalized the way to describe the order of events via **histories**, we can now formalize the mechanism to determine if a **history** is valid. The easier case is for a **sequential history**.

Definition 5.2.1.10 For a concurrent object S , a **sequential specification** of S is a function that either returns true (valid) or false (invalid) for a **sequential history** H .

The harder case is handled via the notion of **linearizable**.

Definition 5.2.1.11 A history H on a concurrent object S is **linearizable** if it has an extension H' and there exists a *sequential history* H_S such that:

1. The **sequential specification** of S accepts H_S .
2. There exists a one-to-one mapping M of a method call $(i, r) \in H'$ to a method call $(i_S, r_S) \in H_S$ with the properties that:
 - i must be the same as i_S except for the timestamp.
 - r must be the same r_S except for the timestamp or r .
3. For any two method calls X and Y in H' ,
$$X \rightarrow_{H'} Y \Rightarrow M(X) \rightarrow_{H_S} M(Y).$$

We consider a history to be valid if it's linearizable.

5.2.1.1 Linearizable SPSC

Our SPSC supports 3 methods:

- enqueue which accepts an input parameter and returns a boolean.
- dequeue which accepts an output parameter and returns a boolean.
- readFront which accepts an output parameter and returns a boolean.

Definition 5.2.1.1.12 An SPSC is **linearizable** if and only if any history produced from the SPSC that does not have overlapping dequeue method calls and overlapping enqueue method calls is *linearizable* according to the following *sequential specification*:

- An enqueue can only be matched by one dequeue.
- A dequeue can only be matched by one enqueue.
- The order of item dequeues is the same as the order of item enqueues.
- An enqueue can only be matched by a later dequeue.
- A dequeue returns `false` when the queue is empty.
- A dequeue returns `true` and matches an enqueue when the queue is not empty.
- An enqueue returns `false` when the queue is full.
- An enqueue would return `true` when the queue is not full and the number of elements should increase by one.
- A read-front would return `false` when the queue is empty.
- A read-front would return `true` and the first element in the queue is read out.

5.2.1.2 Linearizable MPSC queue

An MPSC queue supports 2 methods:

- enqueue which accepts an input parameter and returns a boolean.
- dequeue which accepts an output parameter and returns a boolean.

Definition 5.2.1.2.13 An MPSC queue is **linearizable** if and only if any history produced from the MPSC queue that does not have overlapping dequeue method calls is *linearizable* according to the following *sequential specification*:

- An enqueue can only be matched by one dequeue.
- A dequeue can only be matched by one enqueue.
- The order of item dequeues is the same as the order of item enqueues.
- An enqueue can only be matched by a later dequeue.
- A dequeue returns `false` when the queue is empty.
- A dequeue returns `true` and matches an enqueue when the queue is not empty
- An enqueue that returns `true` will be matched if there are enough dequeues after that.
- An enqueue that returns `false` will never be matched.

5.2.2 ABA-safety

Not every ABA problem is unsafe. We formalize in this section which ABA problem is safe and which is not.

Definition 5.2.2.14 A **modification instruction** on a variable v is an atomic instruction that may change the value of v e.g. a store or a CAS.

Definition 5.2.2.15 A **successful modification instruction** on a variable v is an atomic instruction that changes the value of v e.g. a store or a successful CAS.

Definition 5.2.2.16 A **CAS-sequence** on a variable v is a sequence of instructions of a method m such that:

- The first instruction is a load $v_0 = \text{load}(v)$.
- The last instruction is a CAS($\&v, v_0, v_1$).
- There's no modification instruction on v between the first and the last instruction.

Definition 5.2.2.17 A **successful CAS-sequence** on a variable v is a **CAS-sequence** on v that ends with a successful CAS.

Definition 5.2.2.18 Consider a method m on a concurrent object S . m is said to be **ABA-safe** if and only if for any history of method calls produced from S , we can reorder any successful CAS-sequences inside an invocation of m in the following fashion:

- If a successful CAS-sequence is part of an invocation of m , after reordering, it must still be part of that invocation.
- If a successful CAS-sequence by an invocation of m precedes another by that invocation, after reordering, this ordering is still respected.
- Any successful CAS-sequence by an invocation of m after reordering must not overlap with a successful modification instruction on the same variable.
- After reordering, all method calls' response events on the concurrent object S stay the same.

5.2.3 Performance model

5.3 Theoretical proofs of the distributed SPSC

In this section, we focus on the correctness and progress guarantee of the simple distributed SPSC established in Section 4.2.

5.3.1 Correctness

5.3.1.1 ABA problem

There's no CAS instruction in our simple distributed SPSC, so there's no potential for ABA problem.

5.3.1.2 Memory reclamation

There's no dynamic memory allocation and deallocation in our simple distributed SPSC, so it is memory-safe.

5.3.1.3 Linearizability

We prove that our simple distributed SPSC is linearizable.

Theorem 5.3.1.3.1 (*Linearizability of the simple distributed SPSC*) The distributed SPSC given in Section 4.2 is linearizable.

Proof We claim that the following are the linearization points of our SPSC's methods:

- The linearization point of an `spsc_enqueue` call (Procedure 4) that returns `false` is line 3.
- The linearization point of an `spsc_enqueue` call (Procedure 4) that returns `true` is line 7.
- The linearization point of an `spsc_dequeue` call (Procedure 6) that returns `false` is line 17.
- The linearization point of an `spsc_dequeue` call (Procedure 6) that returns `true` is line 21.
- The linearization point of `spsc_readFronte` call (Procedure 5) that returns `false` is line 10 or line 12 if line 10 is passed.
- The linearization point of `spsc_readFronte` call (Procedure 5) that returns `true` is line 12.
- The linearization point of `spsc_readFrontd` call (Procedure 7) that returns `false` is line 25.
- The linearization point of `spsc_readFrontd` call (Procedure 7) that returns `true` is right after line 25 (or right before line 28 if line 25 is never executed).

We define a total ordering $<$ on the set of completed method calls based on these linearization points: If the linearization point of a method call A is before the linearization point of a method call B , then $A < B$.

If the distributed SPSC is linearizable, $<$ would define a equivalent valid sequential execution order for our SPSC method calls.

A valid sequential execution of SPSC method calls would possess the following characteristics.

An enqueue can only be matched by one dequeue: Each time an `spsc_dequeue` is executed, it advances the `First` index. Because only one dequeue can happen at a time, it's guaranteed that each dequeue proceeds with one unique `First` index. Two dequeues can only dequeue out the same entry in the SPSC's array if their `First` indices are congruent modulo capacity. However, by then, this entry must have been overwritten. Therefore, an enqueue can only be dequeued at most once.

A dequeue can only be matched by one enqueue: This is trivial, as based on how Procedure 6 is defined, a dequeue can only dequeue out at most one value.

The order of item dequeues is the same as the order of item enqueues: To put more precisely, if there are 2 `spsc_enqueues` e_1, e_2 such that $e_1 < e_2$, then either e_2 is unmatched or e_1 matches d_1 and e_2 matches d_2 such that $d_1 < d_2$. If e_2 is unmatched, the statement holds. Suppose e_2 matches d_2 . Because $e_1 < e_2$, based on how Procedure 4 is defined, e_1 corresponds to a value i_1 of `Last` and e_2 corresponds to a value i_2 of `Last` such that $i_1 < i_2$. Based on how Procedure 6 is defined, each time a dequeue happens successfully, `First` would be incremented. Therefore, for e_2 to be matched, e_1 must be matched first because `First` must surpass i_1 before getting to i_2 . In other words, e_1 matches d_1 such that $d_1 < d_2$.

An enqueue can only be matched by a later dequeue: To put more precisely, if an `spsc_enqueue` e matches an `spsc_dequeue` d , then $e < d$. If e hasn't executed its linearization point at line 7, there's no way d 's line 20 can see e 's value. Therefore, d 's linearization point at line 21 must be after e 's linearization point at line 7. Therefore, $e < d$.

A dequeue would return false when the queue is empty: To put more precisely, for an `spsc_dequeue` d , if by d 's linearization point, every successful `spsc_enqueue` e' such that $e' < d$ has been matched by d' such that $d' < d$, then d would be unmatched and return false. By this assumption, any `spsc_enqueue` e that has executed its linearization point at line 7 before d 's line 16 has been matched. Therefore, `First = Last` at line 16, or `First >= Last_buf`, therefore, the if condition at line 16-19 is entered. Also by the assumption, any `spsc_enqueue` e that has executed its linearization point at line 7 before d 's line 18 has been matched. Therefore, `First = Last` at line 18. Then, line 19 is executed and d returns false.

A dequeue would return true and match an enqueue when the queue is not empty: To put more precisely, for an `spsc_dequeue` d , if there exists a successful `spsc_enqueue` e' such that $e' < d$ and has not been matched by a dequeue d' such that $d' < e'$, then d would match some e and return true. By this assumption, some e' must have executed its linearization point at line 7 but is still unmatched by the time d starts. Then, `First < Last`, so d must match some enqueue e and returns true.

An enqueue would return false when the queue is full: To put more precisely, for an `spsc_enqueue` e , if by e 's linearization point, the number of unmatched successful `spsc_enqueue` $e' < e$ by the time e starts equals `Capacity`, then e returns false. By this assumption, any d' that matches e' must satisfy $e < d'$, or d' must execute its synchronization point at line 21 after line 1 and line 4 of e , then e 's line 5 must have executed and return false.

An enqueue would return true when the queue is not full and the number of elements should increase by one: To put more precisely, for an `spsc_enqueue` e , if by e 's linearization point, the number of unmatched successful `spsc_enqueue` $e' < e$ by the time e starts is fewer than `Capacity`, then e returns true. By this assumption, `First < Last` at least until e 's linearization point and because line 7 must be executed, which means the number of elements should increase by one.

A read-front would return false when the queue is empty: To put more precisely, for a read-front r , if by r 's linearization point, every successful `spsc_enqueue` e' such that $e' < r$ has been matched by d' such that $d' < d$, then r would return false. That means any unmatched successful `spsc_enqueue` e must have executed its linearization point at line 7 after r 's, or `First = Tail` before r 's linearization point

- For an enqueueer's read-front, if r doesn't pass line 10, the statement holds. If r passes line 10, by the assumption, r would execute line 14, because r sees that $\text{First} = \text{Tail}$.
- For an dequeuer's read-front, r must enter line 25-27 because $\text{First_buf} = \text{Tail_buf}$, due to from the dequeuer's point of view, $\text{First_buf} = \text{First}$ and $\text{Last_buf} \leq \text{Last}$. Similarly, r must execute line 27 and return false.

A read-front would return true and the first element in the queue is read out: To put more precisely, for a read-front r , if before r 's linearization point, there exists some unmatched successful $\text{spsc_enqueue } e'$ such that $e' < r$, then r would read out the same value as the first d such that $r < d$. By this assumption, any d' that matches some of these successful $\text{spsc_enqueue } e'$ must execute its linearization point at line 21 after r 's linearization point. Therefore, $\text{First} < \text{Last}$ until r 's linearization point.

- For an enqueueer's read-front, r must not execute line 11 and line 14. Therefore, line 15 is executed, and First_buf at this point is the same as First_buf of the first d such that $r < d$, because we have just read it at line 12, and any successful $d' > r$ must execute line 21 after line 15, therefore, First has no chance to be incremented between line 12 and line 15.
- For a dequeuer's read-front, r must not execute line 25-27 and execute line 28 instead. It's trivial that r reads out the same value as the first dequeue d such that $r < d$ because there can only be one dequeuer.

In conclusion, for any completed history of method calls our SPSC can produce, we have defined a way to sequentially order them in a way that conforms to SPSC's sequential specification. By [Definition 5.2.1.1.12](#), our SPSC is linearizable. \square

5.3.2 Progress guarantee

Our simple distributed SPSC is wait-free:

- spsc_dequeue (Procedure 6) does not execute any loops or wait for any other method calls.
- spsc_enqueue (Procedure 4) does not execute any loops or wait for any other method calls.
- spsc_readFront_e (Procedure 5) does not execute any loops or wait for any other method calls.
- spsc_readFront_d (Procedure 7) does not execute any loops or wait for any other method calls.

5.3.3 Performance model

5.4 Theoretical proofs of dLTQueue

5.4.1 Proof-specific notations

The structure of dLTQueue is presented again in Image 3.



Image 3: dLTQueue's structure.

As a reminder, the bottom rectangular nodes are called the **enqueueer nodes** and the circular nodes are called the **tree nodes**. Tree nodes that are attached to an enqueueer node are called **leaf nodes**, otherwise, they are called **internal nodes**. Each **enqueueer node** is hosted on the enqueueer that corresponds to it. The enqueueer nodes accommodate an instance of our distributed SPSC in Section 4.2 and a `Min_timestamp` variable representing the minimum timestamp inside the SPSC. Each **tree node** stores a rank of an enqueueer that's attached to the subtree which roots at the **tree node**.

We will refer propagate_e and propagate_d as propagate if there's no need for discrimination. Similarly, we will sometimes refer to refreshNode_e and refreshNode_d as refreshNode , refreshLeaf_e and refreshLeaf_d as refreshLeaf , $\text{refreshTimestamp}_e$ and $\text{refreshTimestamp}_d$ as refreshTimestamp .

Definition 5.4.1.1 For a tree node n , the rank stored in n at time t is denoted as $\text{rank}(n, t)$.

Definition 5.4.1.2 For an enqueue or a dequeue op, the rank of the enqueueer it affects is denoted as $\text{rank}(\text{op})$.

Definition 5.4.1.3 For an enqueueer whose rank is r , the `Min_timestamp` value stored in its enqueueer node at time t is denoted as $\text{min-ts}(r, t)$. If r is `DUMMY_RANK`, $\text{min-ts}(r, t)$ is `MAX_TIMESTAMP`.

Definition 5.4.1.4 For an enqueueer with rank r , the minimum timestamp among the elements between `First` and `Last` in its SPSC at time t is denoted as $\text{min-spsc-ts}(r, t)$. If r is dummy, $\text{min-spsc-ts}(r, t)$ is `MAX`.

Definition 5.4.1.5 For an enqueue or a dequeue op, the set of nodes that it calls `refreshNode` (Procedure 15 or Procedure 20) or `refreshLeaf` (Procedure 16 or Procedure 21) on is denoted as $path(op)$.

Definition 5.4.1.6 For an enqueue or a dequeue, **timestamp-refresh phase** refer to its execution of line 18-19 in `propagatee` (Procedure 13) or line 71-72 in `propagated` (Procedure 18).

Definition 5.4.1.7 For an enqueue op, and a node $n \in path(op)$, **node- n -refresh phase** refer to its execution of:

- Line 20-21 of `propagatee` (Procedure 13) if n is a leaf node.
- Line 25-26 of `propagatee` (Procedure 13) to refresh n 's rank if n is a non-leaf node.

Definition 5.4.1.8 For a dequeue op, and a node $n \in path(op)$, **node- n -refresh phase** refer to its execution of:

- Line 73-74 of `propagated` (Procedure 18) if n is a leaf node.
- Line 78-79 of `propagated` (Procedure 18) to refresh n 's rank if n is a non-leaf node.

Definition 5.4.1.9 `refreshTimestampe` (Procedure 14) is said to start its **CAS-sequence** if it finishes line 29. `refreshTimestampe` is said to end its **CAS-sequence** if it finishes line 34 or line 36.

Definition 5.4.1.10 `refreshTimestampd` (Procedure 19) is said to start its **CAS-sequence** if it finishes line 83. `refreshTimestampd` is said to end its **CAS-sequence** if it finishes line 88 or line 90.

Definition 5.4.1.11 `refreshNodee` (Procedure 15) is said to start its **CAS-sequence** if it finishes line 38. `refreshNodee` is said to end its **CAS-sequence** if it finishes line 52.

Definition 5.4.1.12 `refreshNoded` (Procedure 20) is said to start its **CAS-sequence** if it finishes line 92. `refreshNoded` is said to end its **CAS-sequence** if it finishes line 106.

Definition 5.4.1.13 `refreshLeafe` (Procedure 16) is said to start its **CAS-sequence** if it finishes line 55. `refreshLeafe` is said to end its **CAS-sequence** if it finishes line 60.

Definition 5.4.1.14 `refreshLeafd` (Procedure 21) is said to start its **CAS-sequence** if it finishes line 109. `refreshLeafd` is said to end its **CAS-sequence** if it finishes line 114.

5.4.2 Correctness

5.4.2.1 ABA problem

We use CAS instructions on:

- Line 34 and line 36 of `refreshTimestampe` (Procedure 14).
- Line 52 of `refreshNodee` (Procedure 15).
- Line 60 of `refreshLeafe` (Procedure 16).
- Line 88 and line 90 of `refreshTimestampd` (Procedure 19).
- Line 106 of `refreshNoded` (Procedure 20).
- Line 114 of `refreshLeafd` (Procedure 21).

Notice that at these locations, we increase the associated version tags of the CAS-ed values. These version tags are 32-bit in size, therefore, practically, ABA problem can't virtually occur. It's safe to assume that there's no ABA problem in dLTQueue.

5.4.2.2 Memory reclamation

Notice that Slotqueue pushes the memory reclamation problem to the underlying SPSC. Because the underlying SPSC is memory-safe, Slotqueue is also memory-safe.

5.4.2.3 Linearizability

Theorem 5.4.2.3.1 In dLTQueue, an enqueue can only match at most one dequeue.

Proof A dequeue indirectly performs a value dequeue through `spsc_dequeue`. Because `spsc_dequeue` can only match one `spsc_enqueue` by another enqueue, the theorem holds. \square

Theorem 5.4.2.3.2 In dLTQueue, a dequeue can only match at most one enqueue.

Proof This is trivial as a dequeue can only read out at most one value, so it can only match at most one enqueue. \square

Theorem 5.4.2.3.3 Only the dequeuer and one enqueuer can operate on an enqueuer node.

Proof This is trivial based on how the algorithm is defined. \square

We immediately obtain the following result.

Corollary 5.4.2.3.4 Only one dequeue operation and one enqueue operation can operate concurrently on an enqueuer node.

Theorem 5.4.2.3.5 The SPSC at an enqueuer node contains items with increasing timestamps.

Proof Each enqueue would FAA the distributed counter (line 14 in Procedure 12) and enqueue into the SPSC an item with the timestamp obtained from that counter. Applying [Corollary 5.4.2.3.4](#), we know that items are enqueued one at a time into the SPSC. Therefore, later items are enqueued by later enqueues, which obtain increasing values by FAA-ing the shared counter. The theorem holds. \square

Theorem 5.4.2.3.6 For an enqueue or a dequeue op, if op modifies an enqueuer node and this enqueuer node is attached to a leaf node l , then $path(op)$ is the set of nodes lying on the path from l to the root node.

Proof This is trivial considering how $propagate_e$ (Procedure 13) and $propagate_d$ (Procedure 18) work. \square

Theorem 5.4.2.3.7 For any time t and a node n , $rank(n, t)$ can only be DUMMY_RANK or the rank of an enqueuer that's attached to the subtree rooted at n .

Proof This is trivial considering how refreshNode_e , refreshNode_d and refreshLeaf_e , refreshLeaf_d works. \square

Theorem 5.4.2.3.8 If an enqueue or a dequeue op begins its **timestamp-refresh phase** at t_0 and finishes at time t_1 , there's always at least one successful call to $\text{refreshTimestamp}_e$ (Procedure 14) or $\text{refreshTimestamp}_d$ (Procedure 19) that affects the enqueuer node corresponding to $\text{rank}(\text{op})$ and this successful call starts and ends its **CAS-sequence** between t_0 and t_1 .

Proof Suppose the interested **timestamp-refresh phase** affects the enqueuer node n .

Notice that the **timestamp-refresh phase** of both enqueue and dequeue consists of at most 2 refreshTimestamp calls affecting n .

If one of the two refreshTimestamp s of the **timestamp-refresh phase** succeeds, then the theorem obviously holds.

Consider the case where both fail.

The first refreshTimestamp fails because there's another refreshTimestamp on n ending its **CAS-sequence** successfully after t_0 but before the end of the first refreshTimestamp 's **CAS-sequence**.

The second refreshTimestamp fails because there's another refreshTimestamp on n ending its **CAS-sequence** successfully after t_0 but before the end of the second refreshTimestamp 's **CAS-sequence**. This another refreshTimestamp must start its **CAS-sequence** after the end of the first successful refreshTimestamp , otherwise, it would overlap with the **CAS-sequence** of the first successful refreshTimestamp , but successful **CAS-sequences** on the same enqueuer node cannot overlap as ABA problem does not occur. In other words, this another refreshTimestamp starts and successfully ends its **CAS-sequence** between t_0 and t_1 .

We have proved the theorem. \square

Theorem 5.4.2.3.9 If an enqueue or a dequeue begins its **node- n -refresh phase** at t_0 and finishes at t_1 , there's always at least one successful refreshNode or refreshLeaf calls affecting n and this successful call starts and ends its **CAS-sequence** between t_0 and t_1 .

Proof This is similar to the above proof. \square

Theorem 5.4.2.3.10 Consider a node n . If within t_0 and t_1 , any dequeue d where $n \in \text{path}(d)$ has finished its **node- n -refresh phase**, then $\min\text{-ts}(\text{rank}(n, t_x), t_y)$ is monotonically decreasing for $t_x, t_y \in [t_0, t_1]$.

Proof We have the assumption that within t_0 and t_1 , all dequeue where $n \in \text{path}(d)$ has finished its **node- n -refresh phase**. Notice that if n satisfies this assumption, any child of n also satisfies this assumption.

We will prove a stronger version of this theorem: Given a node n , time t_0 and t_1 such that within $[t_0, t_1]$, any dequeue d where $n \in \text{path}(d)$ has finished its **node- n -refresh phase**. Consider the last dequeue's **node- n -refresh phase** before t_0 (there maybe none). Take $t_s(n)$ and $t_e(n)$ to be the starting and ending time of the CAS-sequence of the last successful **n -refresh call** during this phase, or if there is none, $t_s(n) = t_e(n) = 0$. Then, $\min\text{-ts}(\text{rank}(n, t_x), t_y)$ is monotonically decreasing for $t_x, t_y \in [t_e(n), t_1]$.

Consider any enqueueer node of rank r that's attached to a satisfied leaf node. For any n' that is a descendant of n , during $t_s(n')$ and t_1 , there's no call to `spsc_dequeue`. Because:

- If an `spsc_dequeue` starts between t_0 and t_1 , the dequeue that calls it hasn't finished its **node- n' -refresh phase**.
- If an `spsc_dequeue` starts between $t_s(n')$ and t_0 , then a dequeue's **node- n' -refresh phase** must start after $t_s(n')$ and before t_0 , but this violates our assumption of $t_s(n')$.

Therefore, there can only be calls to `spsc_enqueue` during $t_s(n')$ and t_1 . Thus, $\min\text{-spsc-ts}(r, t_x)$ can only decrease from `MAX_TIMESTAMP` to some timestamp and remain constant for $t_x \in [t_s(n'), t_1]$. (1)

Similarly, there can be no dequeue that hasn't finished its **timestamp-refresh phase** during $t_s(n')$ and t_1 . Therefore, $\min\text{-ts}(r, t_x)$ can only decrease from `MAX_TIMESTAMP` to some timestamp and remain constant for $t_x \in [t_s(n'), t_1]$. (2)

Consider any satisfied leaf node n_0 . There can't be any dequeue that hasn't finished its **node- n_0 -refresh phase** during $t_e(n_0)$ and t_1 . Therefore, any successful `refreshLeaf` affecting n_0 during $[t_e(n_0), t_1]$ must be called by an enqueue. Because there's no `spsc_dequeue`, this `refreshLeaf` can only set $\text{rank}(n_0, t_x)$ from `DUMMY_RANK` to r and this remains r until t_1 , which is the rank of the enqueueer whose node it's attached to. Therefore, combining with (1), $\min\text{-ts}(\text{rank}(n_0, t_x), t_y)$ is monotonically decreasing for $t_x, t_y \in [t_e(n_0), t_1]$. (3)

Consider any satisfied non-leaf node n' that is a descendant of n . Suppose during $[t_e(n'), t_1]$, we have a sequence of successful **n' -refresh calls** that start their CAS-sequences at $t_{\text{start-0}} < t_{\text{start-1}} < t_{\text{start-2}} < \dots < t_{\text{start-k}}$ and end them at $t_{\text{end-0}} < t_{\text{end-1}} < t_{\text{end-2}} < \dots < t_{\text{end-k}}$. By definition, $t_{\text{end-0}} = t_e(n')$ and $t_{\text{start-0}} = t_s(n')$. We can prove that $t_{\text{end-i}} < t_{\text{start-(i+1)}}$ because successful CAS-sequences cannot overlap.

Due to how `refreshNode` (Procedure 15 and Procedure 20) is defined, for any $k \geq i \geq 1$:

- Suppose $t_{\text{rank-i}}(c)$ is the time `refreshNode` reads the rank stored in the child node c , so $t_{\text{start-i}} \leq t_{\text{rank-i}}(c) \leq t_{\text{end-i}}$.
- Suppose $t_{\text{ts-i}}(c)$ is the time `refreshNode` reads the timestamp stored in the enqueueer with the rank read previously, so $t_{\text{start-i}} \leq t_{\text{ts-i}}(c) \leq t_{\text{end-i}}$.
- There exists a child c_i such that $\text{rank}(n', t_{\text{end-i}}) = \text{rank}(c_i, t_{\text{rank-i}}(c_i))$. (4)
- For every child c of n' ,

$$\begin{aligned} & \min\text{-ts}(\text{rank}(n', t_{\text{end-i}}), t_{\text{ts-i}}(c_i)) \\ & \leq \min\text{-ts}(\text{rank}(c, t_{\text{rank-i}}(c)), t_{\text{ts-i}}(c)). \end{aligned} \quad (5)$$

Suppose the stronger theorem already holds for every child c of n' . (6)

For any $i \geq 1$, we have $t_e(c) \leq t_s(n') \leq t_{start-(i-1)} \leq t_{rank-(i-1)}(c) \leq t_{end-(i-1)} \leq t_{start-i} \leq t_{rank-i}(c) \leq t_1$. Combining with (5), (6), we have for any $k \geq i \geq 1$,

$$\begin{aligned} & \min-ts(rank(n', t_{end-i}), t_{ts-i}(c_i)) \\ & \leq \min-ts(rank(c, t_{rank-i}(c)), t_{ts-i}(c)) \\ & \leq \min-ts(rank(c, t_{rank-(i-1)}(c)), t_{ts-i}(c)). \end{aligned}$$

Choose $c = c_{i-1}$ as in (4). We have for any $k \geq i \geq 1$,

$$\begin{aligned} & \min-ts(rank(n', t_{end-i}), t_{ts-i}(c_i)) \\ & \leq \min-ts(rank(c_{i-1}, t_{rank-(i-1)}(c_{i-1})), t_{ts-i}(c_{i-1})) \\ & = \min-ts(rank(n', t_{end-(i-1)}), t_{ts-i}(c_{i-1})). \end{aligned}$$

Because $t_{ts-i}(c_i) \leq t_{end-i}$ and $t_{ts-i}(c_{i-1}) \geq t_{end-(i-1)}$ and (2), we have for any $k \geq i \geq 1$,

$$\begin{aligned} & \min-ts(rank(n', t_{end-i}), t_{end-i}) \\ & \leq \min-ts(rank(n', t_{end-(i-1)}), t_{end-(i-1)}). (*) \end{aligned}$$

$rank(n', t_x)$ can only change after each successful `refreshNode`, therefore, the sequence of its value is $rank(n', t_{end-0}), rank(n', t_{end-1}), \dots, rank(n', t_{end-k})$. (**)

Note that if `refreshNode` observes that an enqueueer has a `Min_timestamp` of `MAX_TIMESTAMP`, it would never try to CAS n' 's rank to the rank of that enqueueer (line 46 of Procedure 15 and line 100 of Procedure 20). So, if `refreshNode` actually set the rank of n' to some non-DUMMY_RANK value, the corresponding enqueueer must actually has a non-MAX_TIMESTAMP `Min-timestamp` at some point. Due to (2), this is constant up until t_1 . Therefore, $\min-ts(rank(n', t_{end-i}), t)$ is constant for any $t \geq t_{end-i}$ and $k \geq i \geq 1$. $\min-ts(rank(n', t_{end-0}), t)$ is constant for any $t \geq t_{end-0}$ if there's a `refreshNode` before t_0 . If there's no `refreshNode` before t_0 , it is constant to `MAX_TIMESTAMP`. So, $\min-ts(rank(n', t_{end-i}), t)$ is constant for any $t \geq t_{end-i}$ and $k \geq i \geq 0$. (***)

Combining (*), (**), (***), we obtain the stronger version of the theorem. \square

Theorem 5.4.2.3.11 If an enqueue e obtains a timestamp c , finishes at time t_0 and is still **unmatched** at time t_1 , then for any subrange T of $[t_0, t_1]$ that does not overlap with a dequeue, $\min-ts(rank(root, t_r), t_s) \leq c$ for any $t_r, t_s \in T$.

Proof We will prove a stronger version of this theorem: Suppose an enqueue e obtains a timestamp c , finishes at time t_0 and is still **unmatched** at time t_1 . For every $n_i \in path(e)$, n_0 is the leaf node and n_i is the parent of n_{i-1} , $i \geq 1$. If e starts and finishes its **node- n_i -refresh phase** at $t_{start-i}$ and t_{end-i} then for any subrange T of $[t_{end-i}, t_1]$ that does not overlap with a dequeue d where $n_i \in path(d)$ and d hasn't finished its **node n_i refresh phase**, $\min-ts(rank(n_i, t_r), t_s) \leq c$ for any $t_r, t_s \in T$.

If $t_1 < t_0$ then the theorem holds.

Take r_e to be the rank of the enqueueer that performs e .

Suppose e enqueues an item with the timestamp c into the local SPSC at time $t_{enqueue}$. Because it's still unmatched up until t_1 , c is always in the local SPSC during $t_{enqueue}$ to t_1 . Therefore, $min-spsc-ts(r_e, t) \leq c$ for any $t \in [t_{enqueue}, t_1]$. (1)

Suppose e finishes its **timestamp refresh phase** at t_{r-ts} . Because $t_{r-ts} \geq t_{enqueue}$, due to (1), $min-ts(r_e, t) \leq c$ for every $t \in [t_{r-ts}, t_1]$. (2)

Consider the leaf node $n_0 \in path(e)$. Due to (2), $rank(n_0, t)$ is always r_e for any $t \in [t_{end-0}, t_1]$. Also due to (2), $min-ts(rank(n_0, t_r), t_s) \leq c$ for any $t_r, t_s \in [t_{end-0}, t_1]$.

Consider any non-leaf node $n_i \in path(e)$. We can extend any subrange T to the left until we either:

- Reach a dequeue d such that $n_i \in path(d)$ and d has just finished its **node- n_i -refresh phase**.
- Reach t_{end-i} .

Consider one such subrange T_i .

Notice that T_i always starts right after a **node- n_i -refresh phase**. Due to [Theorem 5.4.2.3.9](#), there's always at least one successful refreshNode in this **node- n_i -refresh phase**.

Suppose the stronger version of the theorem already holds for n_{i-1} . That is, if e starts and finishes its **node- n_{i-1} -refresh phase** at $t_{start-(i-1)}$ and $t_{end-(i-1)}$ then for any subrange T of $[t_{end-(i-1)}, t_1]$ that does not overlap with a dequeue d where $n_i \in path(d)$ and d hasn't finished its **node n_{i-1} refresh phase**, $min-ts(rank(n_i, t_r), t_s) \leq c$ for any $t_r, t_s \in T$.

Extend T_i to the left until we either:

- Reach a dequeue d such that $n_i \in path(d)$ and d has just finished its **node- n_{i-1} -refresh phase**.
- Reach $t_{end-(i-1)}$.

Take the resulting range to be T_{i-1} . Obviously, $T_i \subseteq T_{i-1}$.

T_{i-1} satisfies both criteria:

- It's a subrange of $[t_{end-(i-1)}, t_1]$.
- It does not overlap with a dequeue d where $n_i \in path(d)$ and d hasn't finished its **node- n_{i-1} -refresh phase**.

Therefore, $min-ts(rank(n_{i-1}, t_r), t_s) \leq c$ for any $t_r, t_s \in T_{i-1}$.

Consider the last successful refreshNode on n_i ending not after T_i , take $t_{s'}$ and $t_{e'}$ to be the start and end time of this refreshNode's CAS-sequence. Because right at the start of T_i , a **node- n_i -refresh phase** just ends, this refreshNode must be within this **node- n_i -refresh phase**. (4)

This refreshNode's CAS-sequence must be within T_{i-1} . This is because right at the start of T_{i-1} , a **node- n_{i-1} -refresh phase** just ends and $T_{i-1} \supseteq T_i$, T_{i-1} must cover the **node-**

n_i -refresh phase whose end T_i starts from. Combining with (4), $t_{s'} \in T_{i-1}$ and $t_{e'} \in T_i$. (5)

Due to how refreshNode is defined and the fact that n_{i-1} is a child of n_i :

- t_{rank} is the time refreshNode reads the rank stored in n_{i-1} , so that $t_{s'} \leq t_{rank} \leq t_{e'}$. Combining with (5), $t_{rank} \in T_{i-1}$.
- t_{ts} is the time refreshNode reads the timestamp from that rank $t_{s'} \leq t_{ts} \leq t_{e'}$. Combining with (5), $t_{ts} \in T_{i-1}$.
- There exists a time t' , $t_{s'} \leq t' \leq t_{e'}$,
 $min-ts(rank(n_i, t_{e'}), t') \leq min-ts(rank(n_{i-1}, t_{rank}), t_{ts})$. (6)

From (6) and the fact that $t_{rank} \in T_{i-1}$ and $t_{ts} \in T_{i-1}$, $min-ts(rank(n_i, t_{e'}), t') \leq c$.

There shall be no spsc_dequeue starting within $t_{s'}$ till the end of T_i because:

- If there's an spsc_dequeue starting within T_i , then T_i 's assumption is violated.
- If there's an spsc_dequeue starting after $t_{s'}$ but before T_i , its dequeue must finish its **node- n_i -refresh phase** after $t_{s'}$ and before T_i . However, then $t_{e'}$ is no longer the end of the last successful refreshNode on n_i not after T_i .

Because there's no spsc_dequeue starting in this timespan, $min-ts(rank(n_i, t_{e'}), t_{e'}) \leq min-ts(rank(n_i, t_{e'}), t') \leq c$.

If there's no dequeue between $t_{e'}$ and the end of T_i whose **node- n_i -refresh phase** hasn't finished, then by [Theorem 5.4.2.3.10](#), $min-ts(rank(n_i, t_r), t_s)$ is monotonically decreasing for any t_r, t_s starting from $t_{e'}$ till the end of T_i . Therefore, $min-ts(rank(n_i, t_r), t_s) \leq c$ for any $t_r, t_s \in T_i$.

Suppose there's a dequeue whose **node- n_i -refresh phase** is in progress some time between $t_{e'}$ and the end of T_i . By definition, this dequeue must finish it before T_i . Because $t_{e'}$ is the time of the last successful refresh on n_i before T_i , $t_{e'}$ must be within the **node- n_i -refresh phase** of this dequeue and there should be no dequeue after that. By the way $t_{e'}$ is defined, technically, this dequeue has finished its **node- n_i -refresh phase** right at $t_{e'}$. Therefore, similarly, we can apply [Theorem 5.4.2.3.10](#), $min-ts(rank(n_i, t_r), t_s) \leq c$ for any $t_r, t_s \in T_i$.

By induction, we have proved the stronger version of the theorem. Therefore, the theorem directly follows. \square

Corollary 5.4.2.3.12 Suppose *root* is the root tree node. If an enqueue e obtains a timestamp c , finishes at time t_0 and is still **unmatched** at time t_1 , then for any subrange T of $[t_0, t_1]$ that does not overlap with a dequeue, $min-spsc-ts(rank(root, t_r), t_s) \leq c$ for any $t_r, t_s \in T$.

Proof Call t_{start} and t_{end} to be the start and end time of T .

Applying [Theorem 5.4.2.3.11](#), we have that $min-ts(rank(root, t_r), t_s) \leq c$ for any $t_r, t_s \in T$.

Fix t_r so that $rank(root, t_r) = r$. We have that $min-ts(r, t) \leq c$ for any $t \in T$.

$min-ts(r, t)$ can only change due to a successful `refreshTimestamp` on the enqueuer node with rank r . Consider the last successful `refreshTimestamp` on the enqueuer node with rank r not after T . Suppose that `refreshTimestamp` reads out the minimum timestamp of the local SPSC at $t' \leq t_{start}$.

Therefore, $min-ts(r, t_{start}) = min-spvc-ts(r, t') \leq c$.

We will prove that after t' until t_{end} , there's no `spvc_dequeue` on r running.

Suppose the contrary, then this `spvc_dequeue` must be part of a dequeue. By definition, this dequeue must start and end before t_{start} , else it violates the assumption of T . If this `spvc_dequeue` starts after t' , then its `refreshTimestamp` must finish after t' and before t_{start} . But this violates the assumption that the last `refreshTimestamp` not after t_{start} reads out the minimum timestamp at t' .

Therefore, there's no `spvc_dequeue` on r running during $[t', t_{end}]$. Therefore, $min-spvc-ts(r, t)$ remains constant during $[t', t_{end}]$ because it's not `MAX_TIMESTAMP`.

In conclusion, $min-spvc-ts(r, t) \leq c$ for $t \in [t', t_{end}]$.

We have proved the theorem. □

Theorem 5.4.2.3.13 Given a rank r . If within $[t_0, t_1]$, there's no uncompleted enqueues on rank r and all matching dequeues for any completed enqueues on rank r has finished, then $rank(n, t) \neq r$ for every node n and $t \in [t_0, t_1]$.

Proof If n doesn't lie on the path from root to the leaf node that's attached to the enqueuer node with rank r , the theorem obviously holds.

Due to [Corollary 5.4.2.3.4](#), there can only be one enqueue and one dequeue at a time at an enqueuer node with rank r . Therefore, there is a sequential ordering among the enqueues and a sequential ordering within the dequeues. Therefore, it's sensible to talk about the last enqueue before t_0 and the last matched dequeue d before t_0 .

Since all of these dequeues and enqueues work on the same local SPSC and the SPSC is linearizable, d must match the last enqueue. After this dequeue d , the local SPSC is empty.

When d finishes its **timestamp-refresh phase** at $t_{ts} \leq t_0$, due to [Theorem 5.4.2.3.8](#), there's at least one successful `refreshTimestamp` call in this phase. Because the last enqueue has been matched, $min-ts(r, t) = MAX_TIMESTAMP$ for any $t \in [t_{ts}, t_1]$.

Similarly, for a leaf node n_0 , suppose d finishes its **node- n_0 -refresh phase** at $t_{r-0} \geq t_{ts}$, then $rank(n_0, t) = DUMMY_RANK$ for any $t \in [t_{r-0}, t_1]$. (1)

For any non-leaf node $n_i \in path(d)$, when d finishes its **node- n_i -refresh phase** at t_{r-i} , there's at least one successful `refreshNode` call during this phase. Suppose this `refreshNode` call starts and ends at $t_{start-i}$ and t_{end-i} . Suppose $rank(n_{i-1}, t) \neq r$ for $t \in [t_{r-(i-1)}, t_1]$. By the way `refreshNode` is defined after this `refreshNode` call, n_i will store some rank other than r . Because of (1), after this up until t_1 , r never has a chance

to be visible to a refreshNode on node n_i during $[n_{i-1}, t]$. In other words, $rank(n_i, t) \neq r$ for $t \in [t_{r-i}, t_1]$.

By induction, we obtain the theorem. □

Theorem 5.4.2.3.14 In dLTQueue, if an enqueue e precedes another dequeue d , then either:

- d isn't matched.
- d matches e .
- e matches d' and d' precedes d .
- d matches e' and e' precedes e .
- d matches e' and e' overlaps with e .

Proof If d doesn't match anything, the theorem holds. If d matches e , the theorem also holds. Suppose d matches e' , $e' \neq e$.

If e matches d' and d' precedes d , the theorem also holds. Suppose e matches d' such that d precedes d' or is unmatched. (1)

Suppose e obtains a timestamp of c and e' obtains a timestamp of c' .

Because e precedes d and because an MPSC queue does not allow multiple dequeues, from the start of d at t_0 until after line 4 of dequeue (Procedure 17) at t_1 , e has finished and there's no dequeue running that has *actually performed* `spsc_dequeue`. Also by t_0 and t_1 , e is still unmatched due to (1).

Applying [Corollary 5.4.2.3.12](#), $\min\text{-spsc-ts}(rank(root, t_x), t_y) \leq c$ for $t_x, t_y \in [t_0, t_1]$. Therefore, d reads out a rank r such that $\min\text{-spsc-ts}(r, t) \leq c$ for $t \in [t_0, t_1]$. Consequently, d dequeues out a value with a timestamp not greater than c . Because d matches e' , $c' \leq c$. However, $e' \neq e$ so $c' < c$.

This means that e cannot precede e' , because if so, $c < c'$.

Therefore, e' precedes e or overlaps with e . □

Theorem 5.4.2.3.15 In dLTQueue, if d matches e , then either e precedes or overlaps with d .

Proof If d precedes e , none of the local SPSCs can contain an item with the timestamp of e . Therefore, d cannot return an item with a timestamp of e . Thus d cannot match e .

Therefore, e either precedes or overlaps with d . □

Theorem 5.4.2.3.16 In dLTQueue, If a dequeue d precedes another enqueue e , then either:

- d isn't matched.
- d matches e' such that e' precedes or overlaps with e and $e' \neq e$.

Proof If d isn't matched, the theorem holds.

Suppose d matches e' . Applying [Theorem 5.4.2.3.15](#), e' must precede or overlap with d . In other words, d cannot precede e' .

If e precedes or is e' , then d must precede e' , which is contradictory.

Therefore, e' must precede e or overlap with e . □

Theorem 5.4.2.3.17 In dLTQueue, if an enqueue e_0 precedes another enqueue e_1 , then either:

- Both e_0 and e_1 aren't matched.
- e_0 is matched but e_1 is not matched.
- e_0 matches d_0 and e_1 matches d_1 such that d_0 precedes d_1 .

Proof If both e_0 and e_1 aren't matched, the theorem holds.

Suppose e_1 matches d_1 . By [Theorem 5.4.2.3.15](#), either e_1 precedes or overlaps with d_1 .

If e_0 precedes d_1 , applying [Theorem 5.4.2.3.14](#) for d_1 and e_0 :

- d_1 isn't matched, contradictory.
- d_1 matches e_0 , contradictory.
- e_0 matches d_0 and d_0 precedes d_1 , the theorem holds.
- d_1 matches e_1 and e_1 precedes e_0 , contradictory.
- d_1 matches e_1 and e_1 overlaps with e_0 , contradictory.

If d_1 precedes e_0 , applying [Theorem 5.4.2.3.16](#) for d_1 and e_0 :

- d_1 isn't matched, contradictory.
- d_1 matches e_1 and e_1 precedes or overlaps with e_0 , contradictory.

Consider that d_1 overlaps with e_0 , then d_1 must also overlap with e_1 . Call r_1 the rank of the enqueuer that performs e_1 . Call t to be the time d_1 atomically reads the root's rank on line 4 of dequeue (Procedure 17). Because d_1 matches e_1 , d_1 must read out r_1 at t_1 .

If e_1 is the first enqueue of rank r_1 , then t must be after e_1 has started, because otherwise, due to [Theorem 5.4.2.3.13](#), r_1 would not be in *root* before e_1 .

If e_1 is not the first enqueue of rank r_1 , then t must also be after e_1 has started. Suppose the contrary, t is before e_1 has started:

- If there's no uncompleted enqueue of rank r_1 at t and they are all matched by the time t , due to [Theorem 5.4.2.3.13](#), r_1 would not be in *root* at t . Therefore, d_1 cannot read out r_1 , which is contradictory.
- If there's some unmatched enqueue of rank r_1 at t , d_1 will match one of these enqueues instead because:
 - There's only one dequeue at a time, so unmatched enqueues at t remain unmatched until d_1 performs an `spsc_dequeue`.
 - Due to [Corollary 5.4.2.3.4](#), all the enqueues of rank r_1 must finish before another starts. Therefore, there's some unmatched enqueue of rank r_1 finishing before e_1 .
 - The local SPSC of the enqueueer node of rank r_1 is serializable, so d_1 will favor one of these enqueues over e_1 .

Therefore, t must happen after e_1 has started. Right at t , no dequeue is actually modifying the dLTQueue state and e_0 has finished. If e_0 has been matched at t then the theorem holds. If e_0 hasn't been matched at t , applying [Theorem 5.4.2.3.11](#), d_1 will favor e_0 over e_1 , which is a contradiction.

We have proved the theorem. □

Theorem 5.4.2.3.18 In dLTQueue, if a dequeue d_0 precedes another dequeue d_1 , then either:

- d_0 isn't matched.
- d_1 isn't matched.
- d_0 matches e_0 and d_1 matches e_1 such that e_0 precedes or overlaps with e_1 .

Proof If d_0 isn't matched or d_1 isn't matched, the theorem holds.

Suppose d_0 matches e_0 and d_1 matches e_1 .

Suppose the contrary, e_1 precedes e_0 . Applying [Theorem 5.4.2.3.14](#):

- Both e_0 and e_1 aren't matched, which is contradictory.
- e_1 is matched but e_0 is not matched, which contradictory.
- e_1 matches d_1 and e_0 matches d_0 such that d_1 precedes d_0 , which is contradictory.

Therefore, the theorem holds. □

Theorem 5.4.2.3.19 (*Linearizability of dLTQueue*) The dLTQueue algorithm is linearizable.

Proof Suppose some history H produced from the modified dLTQueue algorithm.

If H contains some pending method calls, we can just wait for them to complete (because the algorithm is wait-free, which we will prove later). Therefore, now we consider all H to contain only completed method calls. So, we know that if a dequeue or an enqueue in H is matched or not.

If there are some unmatched enqueues, we can append dequeues sequentially to the end of H until there's no unmatched enqueues. Consider one such H' .

We already have a strict partial order $\rightarrow_{H'}$ on H' .

Because the queue is MPSC, there's already a total order among the dequeues.

We will extend $\rightarrow_{H'}$ to a strict total order $\Rightarrow_{H'}$ on H' as follows:

- If $X \rightarrow_{H'} Y$ then $X \Rightarrow_{H'} Y$. (1)
- If a dequeue d matches e then $e \Rightarrow_{H'} d$. (2)
- If a dequeue d_0 matches e_0 and another dequeue matches e_1 such that $d_0 \Rightarrow_{H'} d_1$ then $e_0 \Rightarrow_{H'} e_1$. (3)
- If a dequeue d overlaps with an enqueue e but does not match e , $d \Rightarrow_{H'} e$. (4)

We will prove that $\Rightarrow_{H'}$ is a strict total order on H' . That is, for every pair of different method calls X and Y , either exactly one of these is true $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ and for any X , $X \not\Rightarrow_{H'} X$.

It's obvious that $X \not\Rightarrow_{H'} X$.

If X and Y are dequeues, because there's a total order among the dequeues, either exactly one of these is true: $X \rightarrow_{H'} Y$ or $Y \rightarrow_{H'} X$. Then due to (1), either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$. Notice that we cannot obtain $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ from (2), (3), or (4).

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. (*)

If X is a dequeue and Y is an enqueue, in this case (3) cannot help us obtain either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$, so we can disregard it.

- If $X \rightarrow_{H'} Y$, then due to (1), $X \Rightarrow_{H'} Y$. By definition, X precedes Y , so (4) cannot apply. Applying [Theorem 5.4.2.3.16](#), either
 - X isn't matched, (2) cannot apply. Therefore, $Y \not\Rightarrow_{H'} X$.
 - X matches e' and $e' \neq Y$. Therefore, X does not match Y , or (2) cannot apply. Therefore, $Y \not\Rightarrow_{H'} X$.

Therefore, in this case, $X \Rightarrow_{H'} Y$ and $Y \not\Rightarrow_{H'} X$.

- If $Y \rightarrow_{H'} X$, then due to (1), $Y \Rightarrow_{H'} X$. By definition, Y precedes X , so (4) cannot apply. Even if (2) applies, it can only help us obtain $Y \Rightarrow_{H'} X$.

Therefore, in this case, $Y \Rightarrow_{H'} X$ and $X \not\Rightarrow_{H'} Y$.

- If X overlaps with Y :
 - If X matches Y , then due to (2), $Y \Rightarrow_{H'} X$. Because X matches Y , (4) cannot apply. Therefore, in this case $Y \Rightarrow_{H'} X$ but $X \not\Rightarrow_{H'} Y$.
 - If X does not match Y , then due to (4), $X \Rightarrow_{H'} Y$. Because X doesn't match Y , (2) cannot apply. Therefore, in this case $X \Rightarrow_{H'} Y$ but $Y \not\Rightarrow_{H'} X$.

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. (**)

If X is an enqueue and Y is an enqueue, in this case (2) and (4) are irrelevant:

- If $X \rightarrow_{H'} Y$, then due to (1), $X \Rightarrow_{H'} Y$. By definition, X precedes Y . Applying [Theorem 5.4.2.3.17](#),
 - Both X and Y aren't matched, then (3) cannot apply. Therefore, in this case, $Y \not\Rightarrow_{H'} X$.
 - X is matched but Y is not matched, then (3) cannot apply. Therefore, in this case, $Y \not\Rightarrow_{H'} X$.
 - X matches d_x and Y matches d_y such that d_x precedes d_y , then (3) applies and we obtain $X \Rightarrow_{H'} Y$.

Therefore, in this case, $X \Rightarrow_{H'} Y$ but $Y \not\Rightarrow_{H'} X$.

- If $Y \rightarrow_{H'} X$, this case is symmetric to the first case. We obtain $Y \Rightarrow_{H'} X$ but $X \not\Rightarrow_{H'} Y$.
- If X overlaps with Y , because in H' , all enqueues are matched, then, X matches d_x and d_y . Because d_x either precedes or succeeds d_y , Applying (3), we obtain either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ and there's no way to obtain the other.

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. (***)

From (*), (**), (***), we have proved that $\Rightarrow_{H'}$ is a strict total ordering that is consistent with $\rightarrow_{H'}$. In other words, we can order method calls in H' in a sequential manner. We will prove that this sequential order is consistent with FIFO semantics:

- An enqueue can only be matched by one dequeue: This follows from [Theorem 5.4.2.3.1](#).
- A dequeue can only be matched by one enqueue: This follows from [Theorem 5.4.2.3.2](#).
- The order of item dequeues is the same as the order of item enqueues: Suppose there are two enqueues e_1, e_2 such that $e_1 \Rightarrow_{H'} e_2$ and suppose they match d_1 and d_2 . Then we have obtained $e_1 \Rightarrow_{H'} e_2$ either because:
 - (3) applies, in this case $d_1 \Rightarrow_{H'} d_2$ is a condition for it to apply.
 - (1) applies, then e_1 precedes e_2 , by [Theorem 5.4.2.3.17](#), d_1 must precede d_2 , thus $d_1 \Rightarrow_{H'} d_2$.

Therefore, if $e_1 \Rightarrow_{H'} e_2$ then $d_1 \Rightarrow_{H'} d_2$.

- An enqueue can only be matched by a later dequeue: Suppose there is an enqueue e matched by d . By (2), obviously $e \Rightarrow_{H'} d$.
 - If the queue is empty, dequeues return false. Suppose a dequeue d such that any $e \Rightarrow_{H'} d$ is all matched by some d' and $d' \Rightarrow_{H'} d$, we will prove that d is unmatched. By [Theorem 5.4.2.3.15](#), d can only match an enqueue e_0 that precedes or overlaps with d .
 - If e_0 precedes d , by our assumption, it's already matched by another dequeue.
 - If e_0 overlaps with d , by our assumption, $d \Rightarrow_{H'} e_0$ because if $e_0 \Rightarrow_{H'} d$, e_0 is already matched by another d' . Then, we can only obtain this because (4) applies, but then d does not match e_0 .

Therefore, d is unmatched.

- A dequeue returns false when the queue is empty: To put more precisely, for a dequeue d , if every successful enqueue e' such that $e' \Rightarrow_{H'} d$ has been matched by d' such that $d' \Rightarrow_{H'} d$, then d would be unmatched and return false. Suppose the contrary, d matches e . By definition, $e \Rightarrow_{H'} d$. This is a contradiction by our assumption.
- A dequeue returns true and matches an enqueue when the queue is not empty: To put more precisely, for a dequeue d , if there exists a successful enqueue e' such that $e' \Rightarrow_{H'} d$ and has not been matched by a dequeue d' such that $d' \Rightarrow_{H'} e'$, then d would be match some e and return true. This follows from [Theorem 5.4.2.3.11](#).
- An enqueue that returns true will be matched if there are enough dequeues after that: Based on how Procedure 12 is defined, when an enqueue returns true, it has successfully execute `spsc_enqueue`. By [Theorem 5.4.2.3.11](#), at some point, it would eventually be matched.
- An enqueue that returns false will never be matched: Based on how Procedure 12 is defined, when an enqueue returns false, the state of `dLTQueue` is not changed, except for the distributed counter. Therefore, it could never be matched.

In conclusion, $\Rightarrow_{H'}$ is a way we can order method calls in H' sequentially that conforms to FIFO semantics. Therefore, we can also order method calls in H sequentially that

conforms to FIFO semantics as we only append dequeues sequentially to the end of H to obtain H' .

We have proved the theorem. □

5.4.3 Progress guarantee

Notice that every loop in dLTQueue is bounded, and no method have to wait for another. Therefore, dLTQueue is wait-free.

5.4.4 Performance model

5.5 Theoretical proofs of Slotqueue

5.5.1 Proof-specific notations

As a refresher, Figure 19 shows the structure of Slotqueue.

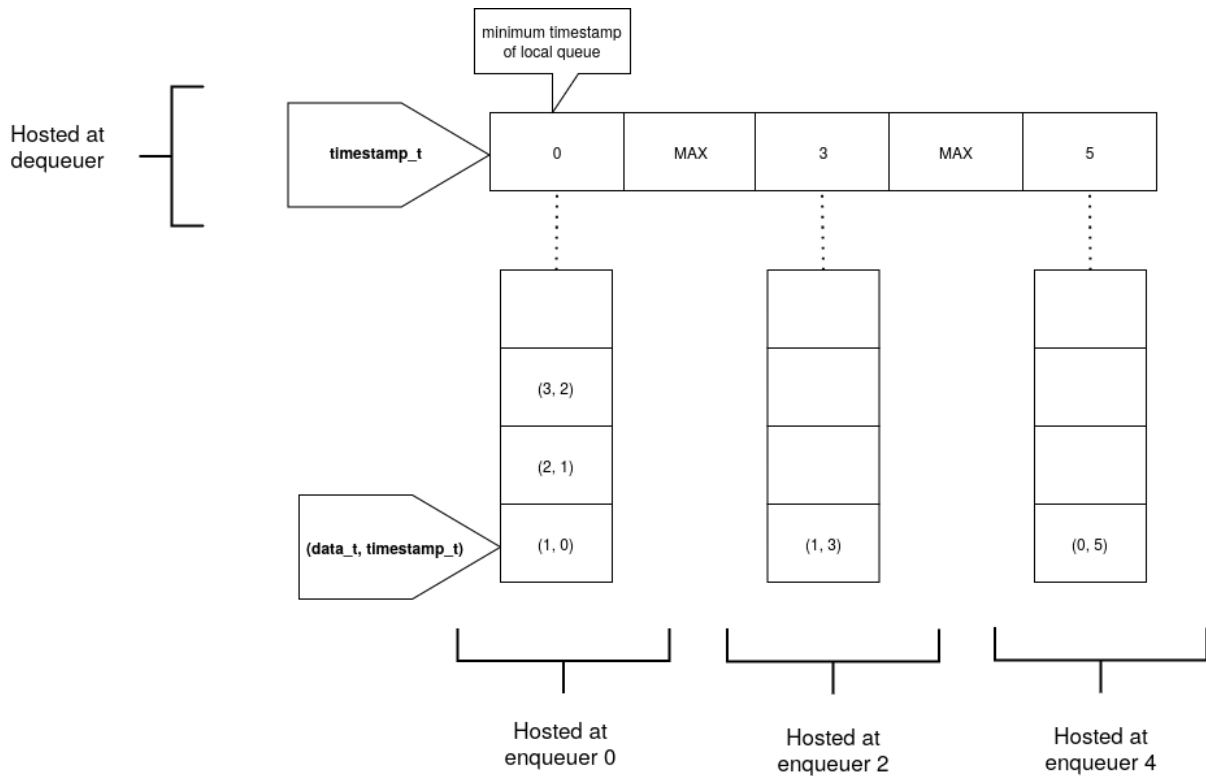


Figure 19: Basic structure of Slotqueue.

Each enqueueur hosts an SPSC that can only accessed by itself and the dequeuer. The dequeuer hosts an array of slots, each slot corresponds to an enqueueur, containing its SPSC's minimum timestamp.

We apply some domain knowledge of Slotqueue algorithm to the definitions introduced in Section 5.2.2.

Definition 5.5.1.1 A **CAS-sequence** on a slot s of an enqueuee that affects s is the sequence of instructions from line 15 to line 20 of its refreshEnqueue (Procedure 25).

Definition 5.5.1.2 A **slot-modification instruction** on a slot s of an enqueue that affects s is line 20 of refreshEnqueue (Procedure 25).

Definition 5.5.1.3 A **CAS-sequence** on a slot s of a dequeue that affects s is the sequence of instructions from line 50 to line 54 of its refreshDequeue (Procedure 28).

Definition 5.5.1.4 A **slot-modification instruction** on a slot s of a dequeue that affects s is line 54 of refreshDequeue (Procedure 28).

Definition 5.5.1.5 A **CAS-sequence** of a dequeue/enqueue is said to **observe a slot value of s_0** if it loads s_0 at line 15 of refreshEnqueue or line 50 of refreshDequeue.

The followings are some other definitions that will be used throughout our proof.

Definition 5.5.1.6 For an enqueue or dequeue op , $rank(op)$ is the rank of the enqueuer whose local SPSC is affected by op .

Definition 5.5.1.7 For an enqueuer whose rank is r , the value stored in its corresponding slot at time t is denoted as $slot(r, t)$.

Definition 5.5.1.8 For an enqueuer with rank r , the minimum timestamp among the elements between First and Last in its local SPSC at time t is denoted as $min-spsc-ts(r, t)$.

Definition 5.5.1.9 For an enqueue, **slot-refresh phase** refer to its execution of line 5-6 of Procedure 24.

Definition 5.5.1.10 For a dequeue, **slot-refresh phase** refer to its execution of line 28-29 of Procedure 26.

Definition 5.5.1.11 For a dequeue, **slot-scan phase** refer to its execution of line 31-47 of Procedure 27.

5.5.2 Correctness

5.5.2.1 ABA problem

Noticeably, we use no scheme to avoid ABA problem in Slotqueue. In actuality, ABA problem does not adversely affect our algorithm's correctness, except in the extreme case that the 64-bit distributed counter overflows, which is unlikely.

We will prove that Slotqueue is ABA-safe, as introduced in Section 5.2.2.

Notice that we only use CASEs on:

- Line 20 of refreshEnqueue (Procedure 25), which is part of an enqueue.
- Line 54 of refreshDequeue (Procedure 28), which is part of a dequeue.

Both CASEs target some slot in the `Slots` array.

Theorem 5.5.2.1.1 (*Concurrent accesses on an SPSC and a slot*) Only one dequeuer and one enqueuer can concurrently modify an SPSC and a slot in the `Slots` array.

Proof This is trivial to prove based on the algorithm's definition. □

Theorem 5.5.2.1.2 (*Monotonicity of SPSC timestamps*) Each SPSC in Slotqueue contains elements with increasing timestamps.

Proof Each enqueue would FAA the distributed counter (line 3 in Procedure 24) and enqueue into the local SPSC an item with the timestamp obtained from the counter. Applying [Theorem 5.5.2.1.1](#), we know that items are enqueued one at a time into the SPSC. Therefore, later items are enqueued by later enqueues, which obtain increasing values by FAA-ing the shared counter. The theorem holds. □

Theorem 5.5.2.1.3 A refreshEnqueue (Procedure 25) can only change a slot to a value other than MAX_TIMESTAMP.

Proof For refreshEnqueue to change the slot's value, the condition on line 18 must be false. Then, new_timestamp must equal to ts, which is not MAX_TIMESTAMP. It's obvious that the CAS on line 20 changes the slot to a value other than MAX_TIMESTAMP. □

Theorem 5.5.2.1.4 (*ABA safety of dequeue*) Assume that the 64-bit distributed counter never overflows, dequeue (Procedure 26) is ABA-safe.

Proof Consider a **successful CAS-sequence** on slot s by a dequeue d . Denote t_d as the value this CAS-sequence observes.

If there's no **successful slot-modification instruction** on slot s by an enqueue e within d 's **successful CAS-sequence**, then this dequeue is ABA-safe.

Suppose the enqueue e executes the **last successful slot-modification instruction** on slot s within d 's **successful CAS-sequence**. Denote t_e to be the value that e sets s (*).

If $t_e \neq t_d$, this CAS-sequence of d cannot be successful, which is a contradiction. Therefore, $t_e = t_d$.

Note that e can only set s to the timestamp of the item it enqueues. That means, e must have enqueued a value with timestamp t_d . However, by definition (*), t_d is read before e executes the CAS, so d cannot observe t_d because e has CAS-ed slot s . This means another process (dequeueer/enqueueer) has seen the value e enqueued and CAS s for e before t_d . By [Theorem 5.5.2.1.1](#), this "another process" must be another dequeueer d' that precedes d because it overlaps with e .

Because d' and d cannot overlap, while e overlaps with both d' and d , e must be the **first** enqueue on s that overlaps with d . Combining with [Theorem 5.5.2.1.1](#) and the fact that e executes the **last successful slot-modification instruction** on slot s within d 's **successful CAS-sequence**, e must be the only enqueue that executes a **successful slot-modification instruction** on s within d 's **successful CAS-sequence**.

During the start of d 's successful CAS-sequence till the end of e , spsc_readFront on the local SPSC must return the same element, because:

- There's no other dequeue running during this time.
- There's no enqueue other than e running.
- The `spsc_enqueue` of e must have completed before the start of d 's successful CAS sequence, because a previous dequeuer d' can see its effect.

Therefore, if we were to move the starting time of d 's successful CAS-sequence right after e has ended, we still retain the output of the program because:

- The CAS sequence only reads two shared values: the `rankth` entry of `Slots` and `spsc_readFront()`, but we have proven that these two values remain the same if we were to move the starting time of d 's successful CAS-sequence this way.
- The CAS sequence does not modify any values except for the last CAS instruction, and the ending time of the CAS sequence is still the same.
- The CAS sequence modifies the `rankth` entry of `Slots` at the CAS but the target value is the same because inputs and shared values are the same in both cases.

We have proved that if we move d 's successful CAS-sequence to start after the **last successful slot-modification instruction** on slot s within d 's **successful CAS-sequence**, we still retain the program's output.

If we apply the reordering for every dequeue, the theorem directly follows. \square

Theorem 5.5.2.1.5 (ABA safety of enqueue) Assume that the 64-bit distributed counter never overflows, enqueue (Procedure 24) is ABA-safe.

Proof Consider a **successful CAS-sequence** on slot s by an enqueue e . Denote t_e as the value this CAS-sequence observes.

If there's no **successful slot-modification instruction** on slot s by a dequeue d within e 's **successful CAS-sequence**, then this enqueue is ABA-safe.

Suppose the dequeue d executes the **last successful slot-modification instruction** on slot s within e 's **successful CAS-sequence**. Denote t_d to be the value that d sets s . If $t_d \neq t_e$, this CAS-sequence of e cannot be successful, which is a contradiction (*).

Therefore, $t_d = t_e$.

If $t_d = t_e = \text{MAX_TIMESTAMP}$, this means e observes a value of `MAX_TIMESTAMP` before d even sets s to `MAX_TIMESTAMP` due to (*). If this `MAX_TIMESTAMP` value is the initialized value of s , it's a contradiction, as s must be non-`MAX_TIMESTAMP` at some point for a dequeue such as d to enter its CAS sequence. If this `MAX_TIMESTAMP` value is set by an enqueue, it's also a contradiction, as `refreshEnqueue` cannot set a slot to `MAX_TIMESTAMP`. Therefore, this `MAX_TIMESTAMP` value is set by a dequeue d' . If $d' \neq d$ then it's a contradiction, because between d' and d , s must be set to be a non-`MAX_TIMESTAMP` value before d can be run, thus, e cannot have observed a value set by d' . Therefore, $d' = d$. But, this means e observes a value set by d , which violates our assumption (*).

Therefore $t_d = t_e = t' \neq \text{MAX_TIMESTAMP}$. e cannot observe the value t' set by d due to our assumption (*). Suppose e observes the value t' from s set by another enqueue/dequeue call other than d .

If this “another call” is a dequeue d' other than d , d' precedes d . By [Theorem 5.5.2.1.2](#), after each dequeue, the front element’s timestamp will be increasing, therefore, d' must have set s to a timestamp smaller than t_d . However, e observes $t_e = t_d$. This is a contradiction.

Therefore, this “another call” is an enqueue e' other than e and e' precedes e . We know that an enqueue only sets s to the timestamp it obtains.

Suppose e' does not overlap with d , then e precedes d . e' can only set s to t' if e' sees that the local SPSC has the front element as the element it enqueues. Due to [Theorem 5.5.2.1.1](#), this means e' must observe a local SPSC with only the element it enqueues. Then, when d executes `readFront`, the item e' enqueues must have been dequeued out already, thus, d cannot set s to t' . This is a contradiction.

Therefore, e' overlaps with d .

Because e' and e cannot overlap, while d overlaps with both e' and e , d must be the *first* dequeue on s that overlaps with e . Combining with [Theorem 5.5.2.1.1](#) and the fact that d executes the **last successful slot-modification instruction** on slot s within e ’s **successful CAS-sequence**, d must be the only dequeue that executes a **successful slot-modification instruction** within e ’s **successful CAS-sequence**.

During the start of e ’s successful CAS-sequence till the end of d , `spsc_readFront` on the local SPSC must return the same element, because:

- There’s no other enqueue running during this time.
- There’s no dequeue other than d running.
- The `spsc_dequeue` of d must have completed before the start of e ’s successful CAS sequence, because a previous enqueuer e' can see its effect.

Therefore, if we were to move the starting time of e ’s successful CAS-sequence right after d has ended, we still retain the output of the program because:

- The CAS sequence only reads two shared values: the `rankth` entry of `slots` and `spsc_readFront()`, but we have proven that these two values remain the same if we were to move the starting time of e ’s successful CAS-sequence this way.
- The CAS sequence does not modify any values except for the last CAS/store instruction, and the ending time of the CAS sequence is still the same.
- The CAS sequence modifies the `rankth` entry of `slots` at the CAS but the target value is the same because inputs and shared values are the same in both cases.

We have proved that if we move e ’s successful CAS-sequence to start after the **last successful slot-modification instruction** on slot s within e ’s **successful CAS-sequence**, we still retain the program’s output.

If we apply the reordering for every enqueue, the theorem directly follows. □

Theorem 5.5.2.1.6 (ABA safety) Assume that the 64-bit distributed counter never overflows, Slot-queue is ABA-safe.

Proof This follows from [Theorem 5.5.2.1.5](#) and [Theorem 5.5.2.1.4](#). □

5.5.2.2 Memory reclamation

5.5.2.3 Linearizability

Theorem 5.5.2.3.7 In Slotqueue, an enqueue can only match at most one dequeue.

Proof A dequeue indirectly performs a value dequeue through `spsc_dequeue`. Because `spsc_dequeue` can only match one `spsc_enqueue` by another enqueue, the theorem holds. □

Theorem 5.5.2.3.8 In Slotqueue, a dequeue can only match at most one enqueue.

Proof This is trivial as a dequeue can only read out at most one value, so it can only match at most one enqueue. □

Theorem 5.5.2.3.9 If an enqueue e begins its **slot-refresh phase** at time t_0 and finishes at time t_1 , there's always at least one successful `refreshEnqueue` that either doesn't execute its **CAS sequence** or starts and ends its **CAS-sequence** between t_0 and t_1 or a successful `refreshDequeue` on $rank(e)$ starting and ending its **CAS-sequence** between t_0 and t_1 .

Proof If one of the two `refreshEnqueues` succeeds, then the theorem obviously holds. Consider the case where both fail.

The first `refreshEnqueue` fails because it tries to execute its **CAS-sequence** but there's another `refreshDequeue` executing its **slot-modification instruction** successfully after t_0 but before the end of the first `refreshEnqueue`'s **CAS-sequence**.

The second `refreshEnqueue` fails because it tries to execute its **CAS-sequence** but there's another `refreshDequeue` executing its **slot-modification instruction** successfully after t_0 but before the end of the second `refreshEnqueue`'s **CAS-sequence**. This another `refreshDequeue` must start its **CAS-sequence** after the end of the first successful `refreshDequeue`, due to [Theorem 5.5.2.1.1](#). In other words, this another `refreshDequeue` starts and successfully ends its **CAS-sequence** between t_0 and t_1 .

We have proved the theorem. □

Theorem 5.5.2.3.10 If a dequeue d begins its **slot-refresh phase** at time t_0 and finishes at time t_1 , there's always at least one successful `refreshEnqueue` or `refreshDequeue` on $rank(d)$ starting and ending its **CAS-sequence** between t_0 and t_1 .

Proof This is similar to the above theorem. □

Theorem 5.5.2.3.11 Given a rank r , if an enqueue e on r that obtains the timestamp c completes at t_0 and is still unmatched by t_1 , then $slot(r, t) \leq c$ for any $t \in [t_0, t_1]$.

Proof Take t' to be the time e 's `spsc_enqueue` takes effect.

At some point after t' , e must enter its **slot-refresh phase**. By [Theorem 5.5.2.3.9](#), there must be a successful refresh call after t' . If this refresh call executes a **CAS-sequence** at $t'' \geq t'$, $t'' \in [t', t_0]$, this **CAS-sequence** must observe the effect of `spsc_enqueue`. Therefore, $\text{slot}(r, t'') \leq c$. If this refresh call doesn't execute a **CAS-sequence**, it must be a `refreshEnqueue` seeing that the front timestamp is different from the enqueued timestamp at t'' , $t'' \in [t', t_0]$. Because e is unmatched up until t_1 and due to [Theorem 5.5.2.1.2](#), $\text{slot}(r, t'') \leq c$.

By the same reasoning as in [Theorem 5.5.2.1.6](#), any successful slot-modification instructions happening after t'' must observe the effect of e 's `spsc_enqueue`. However, because e is never matched between t'' and t_1 , the timestamp c is in the local SPSC the whole timespan $[t'', t_1]$. Therefore, any slot-modification instructions during $[t'', t_1]$ must set the slot's value to some value not greater than c . \square

Theorem 5.5.2.3.12 In Slotqueue, if an enqueue e precedes another dequeue d , then either:

- d isn't matched.
- d matches e .
- e matches d' and d' precedes d .
- d matches e' and e' precedes e .
- d matches e' and e' overlaps with e .

Proof If d doesn't match anything, the theorem holds. If d matches e , the theorem also holds. Suppose d matches e' , $e' \neq e$.

If e matches d' and d' precedes d , the theorem also holds. Suppose e matches d' such that d precedes d' or is unmatched. (1)

Suppose e obtains a timestamp of c and e' obtains a timestamp of c' .

Due to (1), at the time d starts, e has finished but it is still unmatched. By the way Procedure 27 is defined and by [Theorem 5.5.2.3.11](#), d would find a slot that stores a timestamp that is not greater than the one e enqueues. In other word, $c' \leq c$. But $c' \neq c$, then $c' < c$. Therefore, e cannot precede e' , otherwise, $c < c'$.

So, either e' precedes or overlaps with e . The theorem holds. \square

Theorem 5.5.2.3.13 In Slotqueue, if d matches e , then either e precedes or overlaps with d .

Proof If d precedes e , none of the local SPSCs can contain an item with the timestamp of e . Therefore, d cannot return an item with a timestamp of e . Thus d cannot match e .

Therefore, e either precedes or overlaps with d . \square

Theorem 5.5.2.3.14 In Slotqueue, if a dequeue d precedes another enqueue e , then either:

- d isn't matched.
- d matches e' such that e' precedes or overlaps with e and $e' \neq e$.

Proof If d isn't matched, the theorem holds.

Suppose d matches e' . By [Theorem 5.5.2.3.13](#), either e' precedes or overlaps with d . Therefore, $e' \neq e$. Furthermore, e cannot precede e' , because then d would precede e' .

We have proved the theorem. □

Theorem 5.5.2.3.15 If an enqueue e_0 precedes another enqueue e_1 , then either:

- Both e_0 and e_1 aren't matched.
- e_0 is matched but e_1 is not matched.
- e_0 matches d_0 and e_1 matches d_1 such that d_0 precedes d_1 .

Proof If e_1 is not matched, the theorem holds.

Suppose e_1 matches d_1 . By [Theorem 5.5.2.3.13](#), either e_1 precedes or overlaps with d_1 .

Suppose the contrary, e_0 is unmatched or e_0 matches d_0 such that d_1 precedes d_0 , then when d_1 starts, e_0 is still unmatched.

If e_0 and e_1 targets the same rank, it's obvious that d_1 must prioritize e_0 over e_1 . Thus d_1 cannot match e_1 .

If e_0 targets a later rank than e_1 , d_1 cannot find e_1 in the first scan, because the scan is left-to-right, and if it finds e_1 it would later find e_0 that has a lower timestamp. Suppose d_1 finds e_1 in the second scan, that means d_1 finds $e' \neq e_1$ and e' 's timestamp is larger than e_1 's, which is larger than e_0 's. Due to the scan being left-to-right, e' must target a later rank than e_1 . If e' also targets a later rank than e_0 , then in the second scan, d_1 would have prioritized e_0 that has a lower timestamp. Suppose e' targets an earlier rank than e_0 but later than e_1 . Because e_0 's timestamp is larger than e' 's, it must precede or overlap with e . Similarly, e_1 must precede or overlap with e . Because e' targets an earlier rank than e_0 , e_0 's **slot-refresh phase** must finish after e' 's. That means e_1 must start after e' 's **slot-refresh phase**, because e_0 precedes e_1 . But then, e_1 must obtain a timestamp larger than e' , which is a contradiction.

Suppose e_0 targets an earlier rank than e_1 . If d_1 finds e_1 in the first scan, then in the second scan, d_1 would have prioritized e_0 's timestamp. Suppose d_1 finds e_1 in the second scan and during the first scan, it finds $e' \neq e_1$ and e' 's timestamp is larger than e_1 's, which is larger than e_0 's. Due to how the second scan is defined, e' targets a later rank than e_1 , which targets a later rank than e_0 . Because during the second scan, e_0 is not chosen, its **slot-refresh phase** must finish after e' 's. Because e_0 precedes e_1 , e_1 must start after e' 's **slot-refresh phase**, so it must obtain a larger timestamp than e' , which is a contradiction.

Therefore, by contradiction, e_0 must be matched and e_0 matches d_0 such that d_0 precedes d_1 . □

Theorem 5.5.2.3.16 In Slotqueue, if a dequeue d_0 precedes another dequeue d_1 , then either:

- d_0 isn't matched.
- d_1 isn't matched.
- d_0 matches e_0 and d_1 matches e_1 such that e_0 precedes or overlaps with e_1 .

Proof If either d_0 isn't matched or d_1 isn't matched, the theorem holds.

Suppose d_0 matches e_0 and d_1 matches e_1 .

If e_1 precedes e_0 , applying [Theorem 5.5.2.3.15](#), we have e_1 matches d_1 and e_0 matches d_0 such that d_1 precedes d_0 . This is a contradiction.

Therefore, e_0 either precedes or overlaps with e_1 . □

Theorem 5.5.2.3.17 (*Linearizability of Slotqueue*) Slotqueue is linearizable.

Proof Suppose some history H produced from the Slot-queue algorithm.

If H contains some pending method calls, we can just wait for them to complete (because the algorithm is wait-free, which we will prove later). Therefore, now we consider all H to contain only completed method calls. So, we know that if a dequeue or an enqueue in H is matched or not.

If there are some unmatched enqueues, we can append dequeues sequentially to the end of H until there's no unmatched enqueues. Consider one such H' .

We already have a strict partial order $\rightarrow_{H'}$ on H' .

Because the queue is MPSC, there's already a total order among the dequeues.

We will extend $\rightarrow_{H'}$ to a strict total order $\Rightarrow_{H'}$ on H' as follows:

- If $X \rightarrow_{H'} Y$ then $X \Rightarrow_{H'} Y$. (1)
- If a dequeue d matches e then $e \Rightarrow_{H'} d$. (2)
- If a dequeue d_0 matches e_0 and another dequeue matches e_1 such that $d_0 \Rightarrow_{H'} d_1$ then $e_0 \Rightarrow_{H'} e_1$. (3)
- If a dequeue d overlaps with an enqueue e but does not match e , $d \Rightarrow_{H'} e$. (4)

We will prove that $\Rightarrow_{H'}$ is a strict total order on H' . That is, for every pair of different method calls X and Y , either exactly one of these is true $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ and for any X , $X \not\Rightarrow_{H'} X$.

It's obvious that $X \not\Rightarrow_{H'} X$.

If X and Y are dequeues, because there's a total order among the dequeues, either exactly one of these is true: $X \rightarrow_{H'} Y$ or $Y \rightarrow_{H'} X$. Then due to (1), either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$. Notice that we cannot obtain $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ from (2), (3), or (4).

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. (*)

If X is a dequeue and Y is an enqueue, in this case (3) cannot help us obtain either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$, so we can disregard it.

- If $X \rightarrow_{H'} Y$, then due to (1), $X \Rightarrow_{H'} Y$. By definition, X precedes Y , so (4) cannot apply. Applying [Theorem 5.5.2.3.14](#), either
 - X isn't matched, (2) cannot apply. Therefore, $Y \not\Rightarrow_{H'} X$.
 - X matches e' and $e' \neq Y$. Therefore, X does not match Y , or (2) cannot apply. Therefore, $Y \not\Rightarrow_{H'} X$.

Therefore, in this case, $X \Rightarrow_{H'} Y$ and $Y \not\Rightarrow_{H'} X$.

- If $Y \rightarrow_{H'} X$, then due to (1), $Y \Rightarrow_{H'} X$. By definition, Y precedes X , so (4) cannot apply. Even if (2) applies, it can only help us obtain $Y \Rightarrow_{H'} X$. Therefore, in this case, $Y \Rightarrow_{H'} X$ and $X \not\Rightarrow_{H'} Y$.
- If X overlaps with Y :
 - If X matches Y , then due to (2), $Y \Rightarrow_{H'} X$. Because X matches Y , (4) cannot apply. Therefore, in this case $Y \Rightarrow_{H'} X$ but $X \not\Rightarrow_{H'} Y$.
 - If X does not match Y , then due to (4), $X \Rightarrow_{H'} Y$. Because X doesn't match Y , (2) cannot apply. Therefore, in this case $X \Rightarrow_{H'} Y$ but $Y \not\Rightarrow_{H'} X$.

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. ($**$)

If X is an enqueue and Y is an enqueue, in this case (2) and (4) are irrelevant:

- If $X \rightarrow_{H'} Y$, then due to (1), $X \Rightarrow_{H'} Y$. By definition, X precedes Y . Applying [Theorem 5.5.2.3.15](#),
 - Both X and Y aren't matched, then (3) cannot apply. Therefore, in this case, $Y \not\Rightarrow_{H'} X$.
 - X is matched but Y is not matched, then (3) cannot apply. Therefore, in this case, $Y \not\Rightarrow_{H'} X$.
 - X matches d_x and Y matches d_y such that d_x precedes d_y , then (3) applies and we obtain $X \Rightarrow_{H'} Y$.

Therefore, in this case, $X \Rightarrow_{H'} Y$ but $Y \not\Rightarrow_{H'} X$.

- If $Y \rightarrow_{H'} X$, this case is symmetric to the first case. We obtain $Y \Rightarrow_{H'} X$ but $X \not\Rightarrow_{H'} Y$.
- If X overlaps with Y , because in H' , all enqueues are matched, then, X matches d_x and d_y . Because d_x either precedes or succeeds d_y , Applying (3), we obtain either $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ and there's no way to obtain the other.

Therefore, exactly one of $X \Rightarrow_{H'} Y$ or $Y \Rightarrow_{H'} X$ is true. ($***$)

From ($*$), ($**$), ($***$), we have proved that $\Rightarrow_{H'}$ is a strict total ordering that is consistent with $\rightarrow_{H'}$. In other words, we can order method calls in H' in a sequential manner. We will prove that this sequential order is consistent with FIFO semantics:

- An enqueue can only be matched by one dequeue: This follows from [Theorem 5.5.2.3.7](#).
- A dequeue can only be matched by one enqueue: This follows from [Theorem 5.5.2.3.8](#).
- The order of item dequeues is the same as the order of item enqueues: Suppose there are two enqueues e_1, e_2 such that $e_1 \Rightarrow_{H'} e_2$ and suppose they match d_1 and d_2 . Then we have obtained $e_1 \Rightarrow_{H'} e_2$ either because:
 - ▶ (3) applies, in this case $d_1 \Rightarrow_{H'} d_2$ is a condition for it to apply.
 - ▶ (1) applies, then e_1 precedes e_2 , by [Theorem 5.5.2.3.15](#), d_1 must precede d_2 , thus $d_1 \Rightarrow_{H'} d_2$.

Therefore, if $e_1 \Rightarrow_{H'} e_2$ then $d_1 \Rightarrow_{H'} d_2$.

- An enqueue can only be matched by a later dequeue: Suppose there is an enqueue e matched by d . By (2), obviously $e \Rightarrow_{H'} d$.
 - ▶ If the queue is empty, dequeues return false. Suppose a dequeue d such that any $e \Rightarrow_{H'} d$ is all matched by some d' and $d' \Rightarrow_{H'} d$, we will prove that d is unmatched. By [Theorem 5.5.2.3.13](#), d can only match an enqueue e_0 that precedes or overlaps with d .
 - If e_0 precedes d , by our assumption, it's already matched by another dequeue.
 - If e_0 overlaps with d , by our assumption, $d \Rightarrow_{H'} e_0$ because if $e_0 \Rightarrow_{H'} d$, e_0 is already matched by another d' . Then, we can only obtain this because (4) applies, but then d does not match e_0 .

Therefore, d is unmatched.

- A dequeue returns false when the queue is empty: To put more precisely, for a dequeue d , if every successful enqueue e' such that $e' \Rightarrow_{H'} d$ has been matched by d' such that $d' \Rightarrow_{H'} d$, then d would be unmatched and return false. Suppose the contrary, d matches e . By definition, $e \Rightarrow_{H'} d$. This is a contradiction by our assumption.
- A dequeue returns true and matches an enqueue when the queue is not empty: To put more precisely, for a dequeue d , if there exists a successful enqueue e' such that $e' \Rightarrow_{H'} d$ and has not been matched by a dequeue d' such that $d' \Rightarrow_{H'} e'$, then d would be match some e and return true. This follows from [Theorem 5.5.2.3.11](#).
- An enqueue that returns true will be matched if there are enough dequeues after that: Based on how Procedure 24 is defined, when an enqueue returns true, it has successfully execute `spsc_enqueue`. By [Theorem 5.5.2.3.11](#), at some point, it would eventually be matched.
- An enqueue that returns false will never be matched: Based on how Procedure 24 is defined, when an enqueue returns false, the state of Slotqueue is not changed, except for the distributed counter. Therefore, it could never be matched.

In conclusion, $\Rightarrow_{H'}$ is a way we can order method calls in H' sequentially that conforms to FIFO semantics. Therefore, we can also order method calls in H sequentially that

conforms to FIFO semantics as we only append dequeues sequentially to the end of H to obtain H' .

We have proved the theorem. □

5.5.3 Progress guarantee

Notice that every loop in Slotqueue is bounded, and no method have to wait for another. Therefore, Slotqueue is wait-free.

5.5.4 Performance model

Chapter VI Preliminary results

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Chapter VII Conclusion & Future works

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