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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

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 algorithms and data structures need to be devised.

Noticeably, multi-producer single-consumer (MPSC) queue is one of those data structures that are utilized heavily in distributed computing, forming the backbone of many applications. Consequently, an MPSC queue can easily present a performance bottleneck if not designed properly. A desirable distributed MPSC queue should be able to exploit the highly concurrent nature of distributed computing. Currently, in the literature, most distributed data structures are designed from the ground up, completely disregarding the any existing data structures developed in the shared memory area, e.g. [2]. This is partly due to the historical differences between the programming models utilized in these two areas. However, since the introduction of specialized networking hardware RDMA and the improved support of the remote memory access (RMA) programming model in MPI-3, this gap has been bridged. Thus, it has opened up a lot new research ([3]) on reusing the principles in the shared memory literature to distributed computing. One favorable characteristic of concurrent data structures that has been heavily researched in the shared memory literature, which is also equally important in distributed computing, is the property of non-blocking, or in particular, lock-freedom. Lock-freedom guarantees that if some processes suspend or die, other processes can still complete. This provides both progress guarantee and fault-tolerance, especially in distributed computing where nodes can fail any time. Thus, the rest of this document concerns itself with investigating and devising efficient non-blocking distributed MPSC queues. Interestingly, we choose to adapt current MPSC queue algorithms in the shared-memory literature to distributed context, which enables a wealth of accumulated knowledge in this literature.

1.2 Research question

Lock-free MPSC queue and other FIFO variants, such as multi-producer multi-consumer (MPMC), concurrent single-producer single-consumer (SPSC), are 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 algorithms in this domain are highly developed and optimized for performance and scalability. However, most research about MPSC queues or FIFO algorithms in general completely disregard the available state-of-the-art algorithms in the shared memory literature. With the new RDMA networking hardware support and capabilities added to MPI-3 RMA API: lock-free shared-memory

algorithms can 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: Port and adapt potential lock-free shared-memory MSPCs to distributed context using the MPI-3 RMA programming model. If this approach proves to be effective, a huge intellectual reuse of shared-memory MPSC algorithms into the distributed domain is possible. Consequently, there may be no need to develop distributed MPSC queue algorithms from the ground up.

1.3 Thesis overview

1.4 Objective

This thesis aims to:

- Investigate state-of-the-art shared-memory MPSC queues.
- Select and appropriately modify potential MPSC queue algorithms so they can be implemented in popular distributed programming environments.
- Port MPSC queue algorithms using MPI-3 RMA.
- Evaluate various theoretical aspects of ported MPSC queue algorithms: Correctness, progress guarantee, time complexity analysis.
- Benchmark the ported MPSC queue algorithms and compare them with current distributed MPSC queues in the literature.
- Discover distributed-environment-specific optimization opportunities for ported MPSC queue algorithms.

1.5 Scope

- For related works on shared-memory MPSC queues, we only focus on linearizable MPSC queues that support at least lock-free enqueue and dequeue operations.
- Any implementation details, benchmarking and optimizations assume MPI-3 settings.
- For optimizations, we focus on performance-related metrics, e.g. time-complexity (theoretically), throughput (empirically).

1.6 Structure

The rest of this report is structured as follows:

Chapter II discusses the theoretical foundation this thesis is based on and the technical terminology that's heavily utilized in this domain. As mentioned, this thesis investigates state-of-the-art shared-memory MPSC queues. Therefore, we discuss the theory related to the design of concurrent algorithms such as lock-freedom and linearizability, the practical challenges such as the ABA problem and safe memory reclamation problem.

We then explore the utilities offered by C++11 to implement concurrent algorithms and MPI-3 to port shared memory algorithms.

Chapter III surveys the shared-memory literature for state-of-the-art queue algorithms, specifically MPSC queues. We specifically focus on algorithms that have the potential to be ported efficiently to distributed context, such as NUMA-aware or can be made to be NUMA-aware. We then conclude with a comparison of the most potential shared-memory queue algorithms.

Chapter IV documents distributed-versions of potential shared-memory MPSC queue algorithms surveys in Chapter III. It specifically presents our adaptation efforts of existing algorithms in the shared-memory literature to make their distributed implementations feasible.

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 introduces our benchmarking setup, including metrics, environments, benchmark/microbenchmark suites and conducting methods. We aim to demonstrate some preliminary results on how well ported shared-memory MPSC queues can compare to existing distributed MPSC 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 treatments to achieve high performance. One specific challenge is that this type of applications is hard to model in traditional MPI APIs. The introduction of MPI RMA (remote memory access) in MPI-2 and its improvement in MPI-3 has significantly improved MPI's capability to express irregular applications comfortably.

2.1.1 The pattern (?)

2.1.2 Multiple-producer, single-consumer (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 in parallel. Concurrent queues have many important applications, namely event handling, scheduling, etc. One class of concurrent FIFOs is 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 queue data structures (MPMC queues) 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. Thus, MPSC queues can see many use cases like MPMCs while being easily scalable and performant.

2.2 Correctness condition of concurrent algorithms - Linearizability

Correctness of concurrent algorithms is hard to define, especially when it comes to 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, linearizability 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.

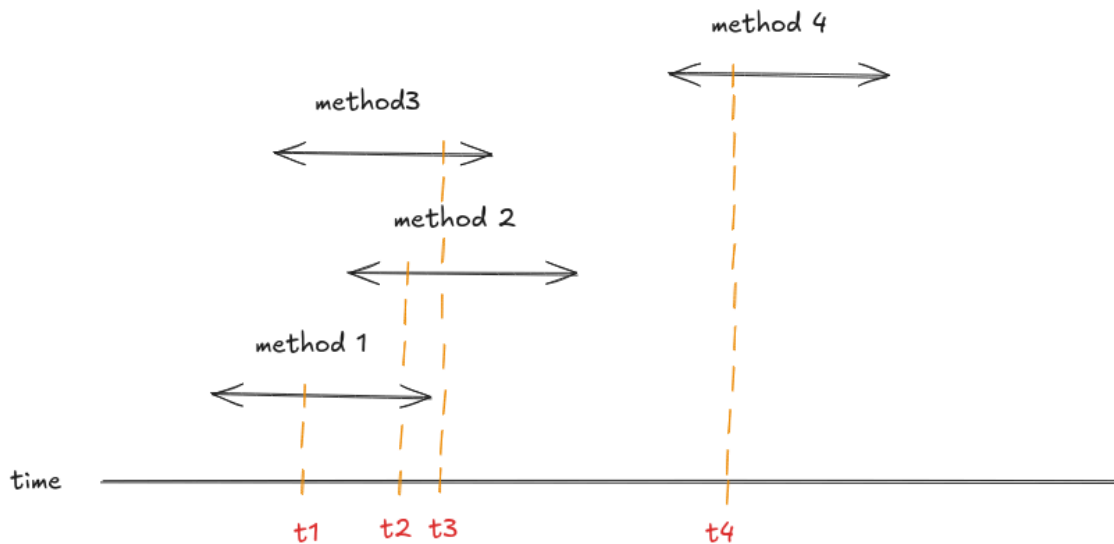


Figure 1: Linearization points of method 1, method 2, method 3, method 4 happens at $t_1 < t_2 < t_3 < t_4$, therefore, their effects will be observed in this order as if we call method 1, method 2, method 3, method 4 sequentially

2.3 Progress guarantee of concurrent algorithms

2.3.1 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. Lock-based 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 halt the whole program. 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 lock-based algorithms 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. These algorithms are said to be **blocking**. 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 to especially interesting subclasses of non-blocking algorithms. Unlike lock-based algorithms, they provide **progress guarantee**.

2.3.2 Non-blocking algorithms

2.3.2.1 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. 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.

2.3.2.2 Wait-free algorithms

Wait-freedom is a stronger progress guarantee than lock-freedom. 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 bounded time. Wait-freedom is useful to have 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.

2.4 Popular atomic instructions in designing non-blocking algorithms

2.4.1 Fetch-and-add (FAA)

2.4.2 Compare-and-swap (CAS)

2.4.3 Load-linked/Store-conditional (LL/SC)

2.5 Common issues when designing non-blocking algorithms

2.5.1 ABA problem

In implementing concurrent non-blocking algorithms, hardware atomic instructions are utilized to achieve linearizability. The most popular atomic operation instruction is compare-and-swap (CAS). The reason for its popularity is (1) CAS is a **universal atomic instruction** - it has the **consensus number** of ∞ - which means it's the most powerful atomic instruction [10] (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 semantic 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. Concurrent algorithms often utilize CAS as follows:

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 `new value` because `old value` is no longer there, so its associated resources are not valid.

This scheme is susceptible to the notorious 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.

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 the memory location to the same value. Some notable schemes are **monotonic version tag** ([6]) and **hazard pointer** ([11]).

2.5.2 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. Solutions 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. Examples include using a reference counter to count the number of processes holding a reference to some memory and **hazard pointer** [11] to announce to other processes that some memory is not to be freed.

2.6 MPI-3 - A popular distributed programming library interface

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].

2.6.1 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]. 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.

2.6.2 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.

2.6.3 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`.

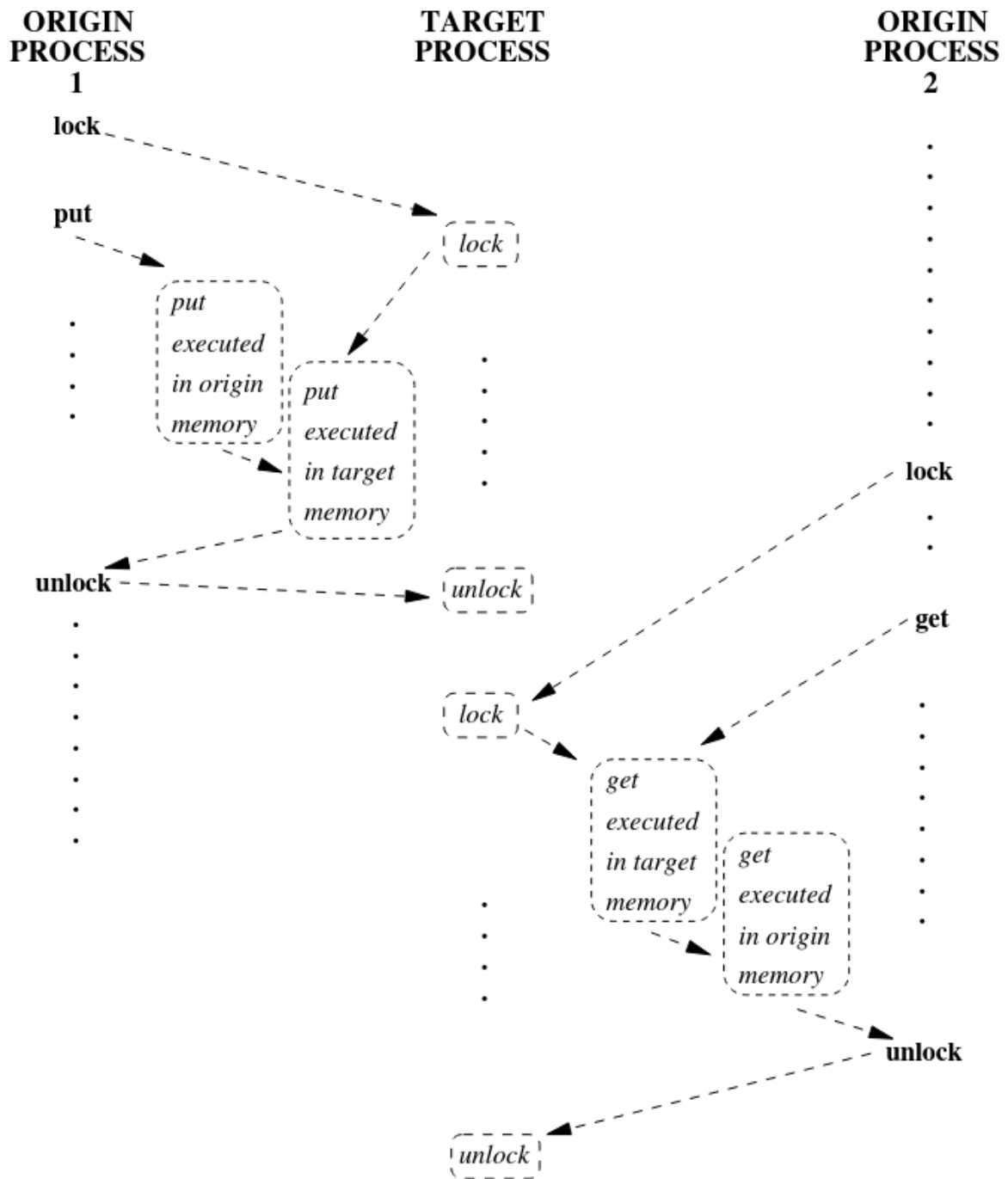


Figure 2: An illustration of passive target communication. Dashed arrows represent synchronization (source: [1])

2.7 Pure MPI - A porting approach of shared memory 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 1: An example snippet showcasing our synchronization approach in MPI RMA

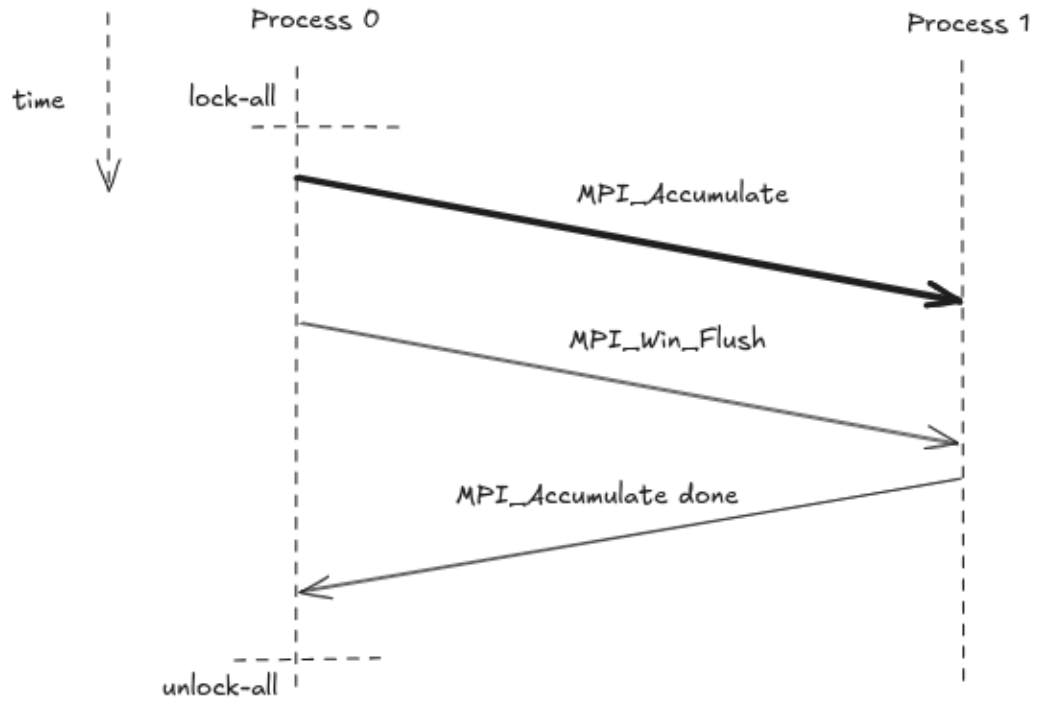


Figure 3: An illustration of our synchronization approach in MPI RMA

Chapter III Related works

There exists numerous research into the design of lock-free shared memory MPMC 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	DQueue	WR-LQueue	Jiffy
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	Un- bounded	Un- bounded	Un- bounded	Un- bounded

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

LTQueue [7] is the earliest wait-free shared memory MPSC queue to our knowledge. This algorithm is wait-free with $O(\log n)$ time complexity for both enqueues and dequeues, with n being the number of enqueueers. Their main idea is to split the MPSC queue among the enqueueers so that each enqueueer maintains a local SPSC data structure, which is only shared with the dequeuer. This improves the MPSC queue's scalability as multiple enqueues can complete the same time. The enqueueers shared a distributed counter and use it to label each item in their local SPSC with a specific timestamp. The timestamps are organized into nodes of a min-heap-like tree so that the dequeuer can look at the root of tree to determine which local SPSC to dequeue next. The min-heap property of the tree is preserved by a novel wait-free timestamp-refreshing operation. Memory reclamation becomes trivial as each MPSC queue entry is only shared by one enqueueer and one dequeuer in the local SPSC. The algorithm avoids ABA problem by utilizing load-link/store-conditional (LL/SC). This, on the other hand, presents a challenge in directly porting LTQueue as LL/SC is not widely available as the more popular CAS instruction.

DQueue [13] focuses on optimizing performance. It aims to be cache-friendly by having each enqueueer batches their updates in a local buffer to decrease cache misses. It also

try to replace expensive atomic instructions such as CAS as many as possible. The MPSC queue is represented as a linked list of segments (which is an array). To enqueue, the enqueuer reserves a slot in the segment list and enqueues the value into the local buffer. If the local buffer is full, the enqueuer flushes the buffer and writes it onto every reserved slot in the segment list. The producer dequeues the values in the segment list in order, upon encountering a reserved but empty slot, it helps all enqueueers flush their local buffers. For memory reclamation, DQueue utilized a dedicated garbage collection thread that reclaims all fully dequeued segments. However, their algorithm is flawed and a segment maybe freed while some process is holding a reference to it.

WRLQueue [14] is a lock-free MPSC queue for embedded real-time system. Its main purpose is to avoid excessive modification of storage space. WRLQueue is simply a pair of buffer, one is worked on by multiple enqueueers and the other is work on by the dequeuer. The enqueueers batch their enqueues and write multiple elements onto the buffer once at a time. The dequeuer upon invocation will swap its buffer with the enqueueer's buffers to dequeue from it. However, this requires the dequeuer to wait for all enqueue operations to complete in their buffer. If an enqueue suspends or dies, the dequeuer will have to wait forever, this clearly violates the property of non-blocking.

Jiffy [8] is a fast and memory-efficient wait-free MPSC queue by avoiding excessive allocation of memory. Like DQueue, Jiffy represents the queue as a linked list of segments. Each enqueue reserves a slot in the segment, extends the linked-list as appropriately, writes the value into the slot and sets a per-slot flag to indicate that the slot is ready to be dequeued. To dequeue, the dequeuer repeatedly scan all the slots to find the first-ready-to-be-dequeue slot. Jiffy shows significant good memory usage and throughput compared to other previous state-of-the-art MPMCs.

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 queue	dLTQueue	Slotqueue
Correctness	Linearizable	Linearizable
Progress guarantee of dequeue	Wait-free	Wait-free
Progress guarantee of enqueue	Wait-free	Wait-free
Worst-case time complexity of dequeue	$O(\log n) R + O(\log n) A$	constant $R + O(n) A$
Worst-case time complexity of enqueue	$O(\log n) R + O(\log n) A$	constant $R + \text{constant } A$
ABA solution	Unique timestamp	No harmful ABA problem
Memory reclamation	Custom scheme	Custom scheme
Number of elements	Unbounded	Unbounded

Table 3: 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 τ 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 position `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.

`T fetch_and_add_sync(remote<T> dest, T 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 undequed 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 enqueuer are given as follows.

Procedure 2: `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 awrite_sync(Data, Last_buf % Capacity, &v)
7 awrite_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 indicating by the difference 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 3: `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_readFronte` 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 enqueueer 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 4: 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 5: bool spsc_readFront_d(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_readFront_d 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 1.

We differentiate between 2 types of nodes: **enqueuer nodes** (represented as the rectangular boxes at the bottom of Image 1) and normal **tree nodes** (represented as the circular boxes in Image 1).

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

Each tree node stores the rank of an enqueuer process. This rank corresponds to the enqueuer node with the minimum timestamp among the node's children's ranks. The tree node that's attached to an enqueuer 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 enqueuer node is set to `MAX_TIMESTAMP` and the corresponding leaf node's rank is set to `DUMMY_RANK`.

Placement-wise:

- The **enqueuer nodes** are hosted at the corresponding **enqueuer**.
- 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

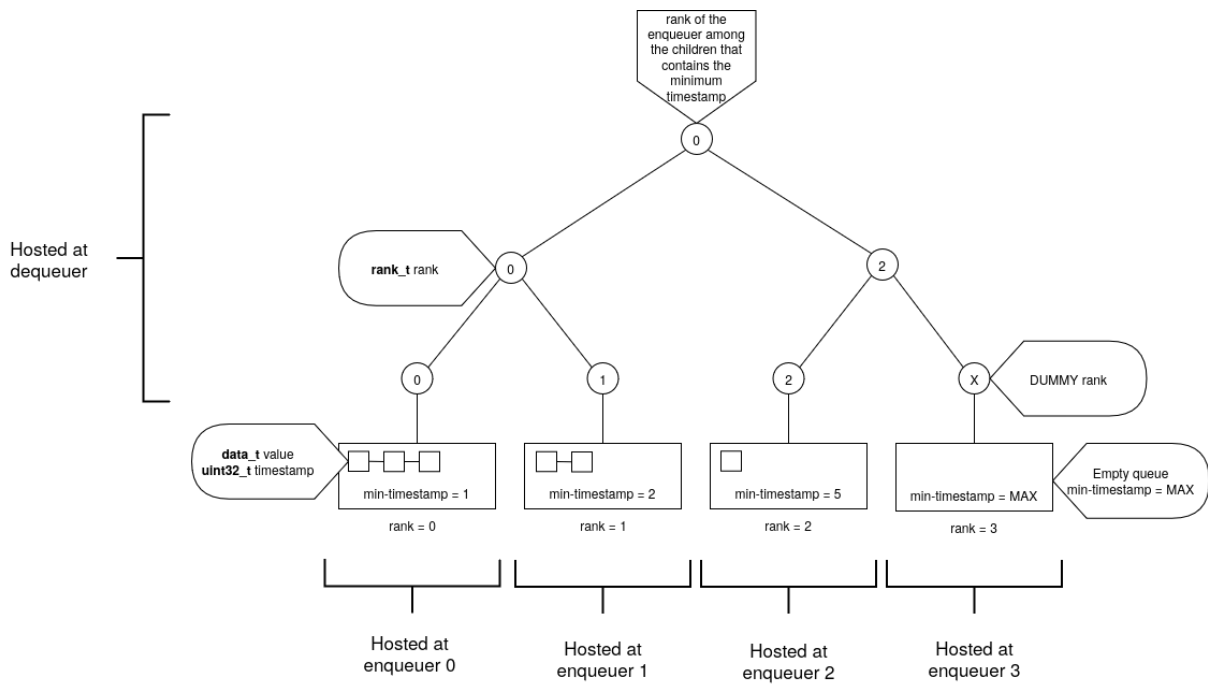


Image 1: dLTQueue's structure

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

`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 7),

children() (Procedure 8), leafNodeIndex() (Procedure 9) 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 6: 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 7: 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 8: `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 9: `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 1.

The followings are the enqueue procedures.

Procedure 10: `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 11: 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 enqueueer node (line 18-19), refreshing of the enqueueer'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 enqueueer node and the root node.

Procedure 12: 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 enqueueer node. It simply looks at the front of the local SPSC (line 310 and CAS Min_timestamp accordingly (line 33-36).

Procedure 13: `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 14: `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 15: `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 16: `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 17: `bool refreshTimestampd(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 `refreshTimestampd` procedure is similar to `refreshTimestampe`, with appropriate changes to accommodate the dequeuer.

Procedure 18: `bool refreshNoded(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 19: $\text{bool refreshLeaf}_d(\text{uint32_t enqueue_rank})$

```

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, that is 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 4.

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.

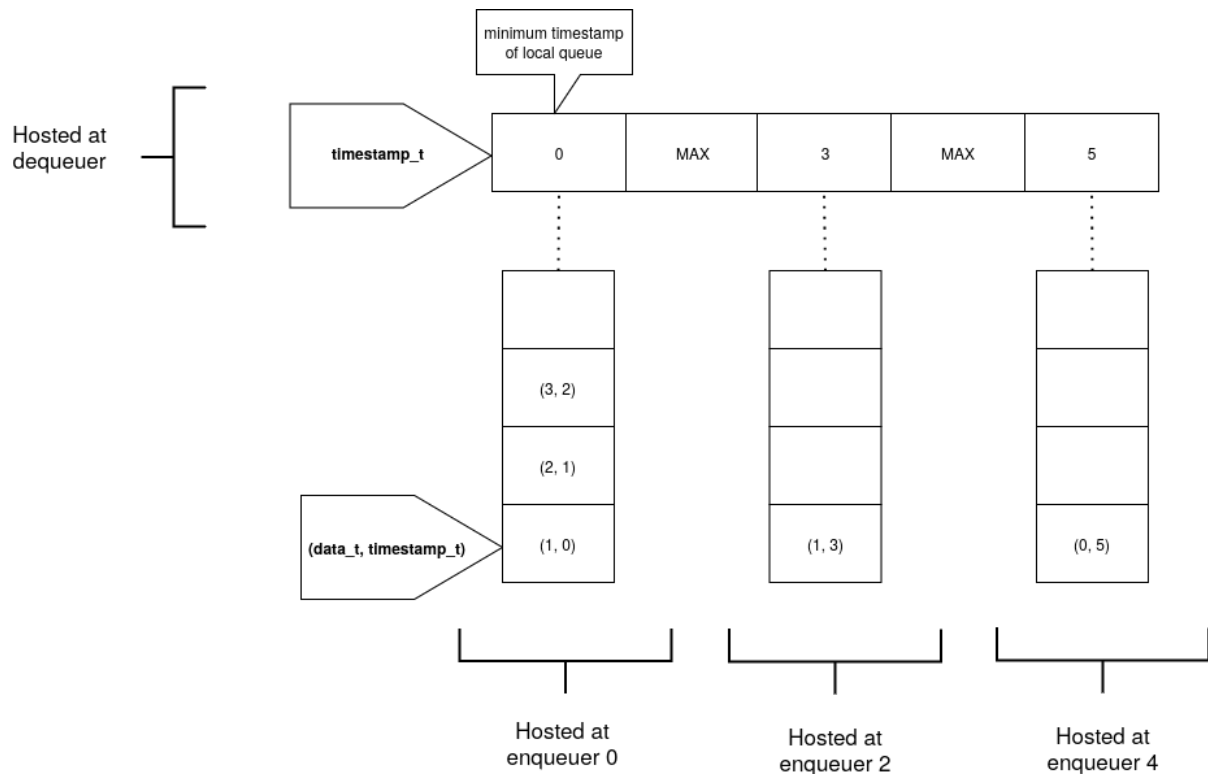


Figure 4: Basic structure of Slotqueue

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 20: 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 21: 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 22: 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 enqueueer tries to refresh its slot's timestamp.

Procedure 23: `bool refreshEnqueue(timestamp_t ts)`

```
8 enqueueer_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, enqueueer_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, enqueueer_order,
20     old_timestamp,
        new_timestamp)
```

`refreshEnqueue`'s responsibility is to refresh the timestamp stores in the enqueueer'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 24: 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 25: 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 26: 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 System model

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 2) that returns `false` is line 3.
- The linearization point of an `spsc_enqueue` call (Procedure 2) that returns `true` is line 7.
- The linearization point of an `spsc_dequeue` call (Procedure 4) that returns `false` is line 17.
- The linearization point of an `spsc_dequeue` call (Procedure 4) that returns `true` is line 21.
- The linearization point of `spsc_readFronte` call (Procedure 3) that returns `false` is line 10 or line 12 if line 10 is passed.
- The linearization point of `spsc_readFronte` call (Procedure 3) that returns `true` is line 12.
- The linearization point of `spsc_readFrontd` call (Procedure 5) that returns `false` is line 25.
- The linearization point of `spsc_readFrontd` call (Procedure 5) 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 4 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 2 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 4 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 4) does not execute any loops or wait for any other method calls.
- spsc_enqueue (Procedure 2) does not execute any loops or wait for any other method calls.
- spsc_readFront_e (Procedure 3) does not execute any loops or wait for any other method calls.
- spsc_readFront_d (Procedure 5) 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 2.



Image 2: 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 13 or Procedure 18) or `refreshLeaf` (Procedure 14 or Procedure 19) 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 11) or line 71-72 in `propagated` (Procedure 16).

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 11) if n is a leaf node.
- Line 25-26 of `propagatee` (Procedure 11) 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 16) if n is a leaf node.
- Line 78-79 of `propagated` (Procedure 16) to refresh n 's rank if n is a non-leaf node.

Definition 5.4.1.9 `refreshTimestampe` (Procedure 12) 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 17) 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 13) 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 18) 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 14) 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 19) 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 12).
- Line 52 of `refreshNodee` (Procedure 13).
- Line 60 of `refreshLeafe` (Procedure 14).
- Line 88 and line 90 of `refreshTimestampd` (Procedure 17).
- Line 106 of `refreshNoded` (Procedure 18).
- Line 114 of `refreshLeafe` (Procedure 19).

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 10) 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 `propagatee` (Procedure 11) and `propagated` (Procedure 16) 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 12) or $\text{refreshTimestamp}_d$ (Procedure 17) 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 13 and Procedure 18) 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 13 and line 100 of Procedure 18). 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 constantlt 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 $t \in T$.

$\text{min-ts}(r, t)$ 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 $t' \leq t_{\text{start}}$.

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

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

Suppose the contrary, then this `spsc_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 `spsc_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 `spsc_dequeue` on r running during $[t', t_{\text{end}}]$. Therefore, $\text{min-spsc-ts}(r, t)$ remains constant during $[t', t_{\text{end}}]$ because it's not `MAX_TIMESTAMP`.

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

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 $\text{rank}(n, t) \neq r$ 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, $\text{min-ts}(r, t) = \text{MAX_TIMESTAMP}$ 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 $\text{rank}(n_0, t) = \text{DUMMY_RANK}$ for any $t \in [t_{r-0}, t_1]$. (1)

For any non-leaf node $n_i \in \text{path}(d)$, 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 $t_{\text{start}-i}$ and $t_{\text{end}-i}$. Suppose $\text{rank}(n_{i-1}, t) \neq r$ 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. \square

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 15) 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 . \square

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 . \square

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 enqueue that performs e_1 . Call t to be the time d_1 atomically reads the root's rank on line 4 of dequeue (Procedure 15). 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 10 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 10 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 5 shows the structure of Slotqueue.

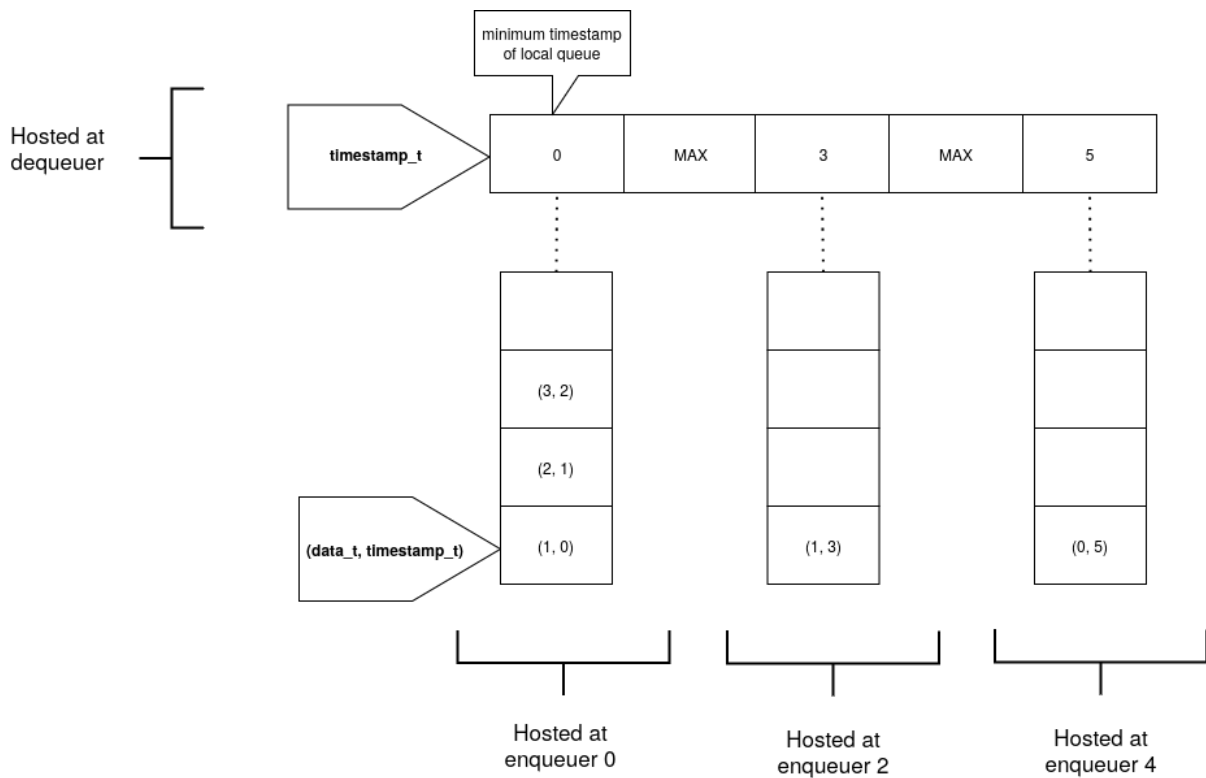


Figure 5: Basic structure of Slotqueue

Each enqueueer hosts an SPSC that can only accessed by itself and the dequeuer. The dequeuer hosts an array of slots, each slot corresponds to an enqueueer, 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 23).

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 23).

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 26).

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 26).

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 22.

Definition 5.5.1.10 For a dequeue, **slot-refresh phase** refer to its execution of line 28-29 of Procedure 24.

Definition 5.5.1.11 For a dequeue, **slot-scan phase** refer to its execution of line 31-47 of Procedure 25.

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 23), which is part of an enqueue.
- Line 54 of refreshDequeue (Procedure 26), 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 22) 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 23) 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 24) 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 22) 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 25 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 22 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 22 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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