Lower Bounds for Restricted-Use Objects
(Extended Abstract)

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Abstract

Concurrent objects play a key role in the design of applications for multi-core architectures, making it imperative to precisely understand their complexity requirements. For some objects, it is known that implementations can be significantly more efficient when their usage is restricted. However, apart from the specific restriction of one-shot implementations, where each process may apply only a single operation to the object, very little is known about the complexities of objects under general restrictions.

This paper draws a more complete picture by defining a large class of objects for which an operation applied to the object can be “perturbed” L consecutive times, and proving lower bounds on the time and space complexity of deterministic implementations of such objects. This class includes bounded-value max registers, limited-use approximate and exact counters, and limited-use collect and compare-and-swap objects; L depends on the number of times the object can be accessed or the maximum value it can support.

For implementations that use only historyless primitives, we prove lower bounds of \( \Omega(\min(\log L, n)) \) on the worst-case step complexity of an operation, where \( n \) is the number of processes; we also prove lower bounds of \( \Omega(\min(L, n)) \) on the space complexity of these objects. When arbitrary primitives can be used, we prove that either some operation incurs \( \Omega(\min(\log L, n)) \) memory stalls or some operation performs \( \Omega(\min(\log L, n)) \) steps.

In addition to these deterministic lower bounds, the paper establishes a lower bound on the expected step complexity of restricted-use randomized approximate counting in a weak oblivious adversary model.
1 Introduction

With multi-core and multi-processor systems now prevalent, there is growing need to gain better understanding of concurrent objects and, specifically, to establish lower bounds on the cost of implementing them. An important general class of concurrent objects, defined by Jayanti, Tan and Toueg [15], are perturbable objects, including widely-used objects, such as counters, max registers, compare-and-swap, single-writer snapshot and fetch-and-add.

Lower bounds are known for long-lived implementations of perturbable objects, where processes apply an unbounded number of operations to the object. For example, Jayanti et al. [15] consider obstruction-free implementations of perturbable objects from historyless primitives, such as read, write, test-and-set and swap. They prove that such implementations require $\Omega(n)$ space and that the worst-case step-complexity of the operations they support is $\Omega(n)$, where $n$ is the number of processes sharing the object.

In some applications, however, objects are used in a restricted manner. For example, there might be a bound on the total number of operations applied on the object, or a bound on the values that the object needs to support. When an object is designed to allow only restricted use, it is sometimes possible to construct more efficient implementations than for the general case.

Indeed, Aspnes, Attiya and Censor-Hillel [3] showed that at least some restricted-use perturbable objects admit implementations that “beat” the lower bound of [15]. For example, a max register can do a write of $v$ in $O(\min(\log v, n))$ steps, while a counter limited to $m$ increments can do each increment in $O(\min(\log^2 m, n))$ steps. Such a restricted-use counter leads to a randomized consensus algorithm with $O(n)$ individual step complexity [4], while restricted-use counters and max registers are used in a mutual exclusion algorithm with sub-logarithmic amortized work [7].

This raises the natural question of determining lower bounds on the complexity of restricted-use objects. The proof of Jayanti et al. [15] breaks for restricted-use objects because the executions constructed by these proofs exceed the restrictions on these objects.

For the specific restriction of one-time object implementations, where each process applies exactly one operation to the object, there are lower bounds which are logarithmic in the number of processes, for specific objects [1,2,6] and generic perturbable objects [14]. Yet, these techniques yield bounds that are far from the upper bounds, e.g., when the object can be perturbed a super-polynomial number of times.

This paper provides a more complete picture of the cost of implementing restricted-use objects by studying the middle ground. We give time and space lower bounds for implementations of objects that are only required to work under restricted usage, for general families of restrictions.

We define the notion of $L$-perturbable objects that strictly generalizes classical perturbability; specific examples are bounded-value max registers, limited-use approximate and exact counters, and limited-use compare-and-swap and collect objects.\(^1\) $L$, the perturbation bound, depends on the number of times the object can be accessed or the maximum value it can support (see Table 1).

For $L$-perturbable objects, we show lower bounds on the step and space complexity of obstruction-free deterministic implementations from historyless primitives. The step complexity lower bound is $\Omega(\min(\log L, n))$, and its proof employs a technique that we call backtracking covering, introduced by Fich, Hendler and Shavit in [11] and later used in [5]. The space complexity lower bound is $\Omega(\min(L, n))$.

\(^1\)A single-writer snapshot object is also a collect object (the converse is, in general, false). Therefore, our lower bounds for the collect object also hold for the single-writer snapshot object.
We also consider implementations that can apply arbitrary primitives and not only historyless primitives, and use the memory stalls measure [8] to quantify the contention incurred by such implementations. We use backtracking covering to prove that either an implementation’s worst-case operation step complexity is $\Omega(\min(\log L, n))$ or some operation incurs $\Omega(\min(\log \log L, \log \log \log L))$ stalls.

In addition to our deterministic lower bounds, we establish a lower bound of $\Omega(\log m / \log \log m)$ on the expected step complexity of randomized $m$-valued $c$-accurate counters, a particularly weak class of counters that allow a multiplicative error of factor at most $c$. Our lower bound employs Yao’s Principle [16] and assumes a weak oblivious adversary.

Table 1 summarizes the lower bounds for specific $L$-perturbable objects.

Aspnes et al. [3] prove lower bounds on obstruction-free implementations of max registers and approximate counters from historyless primitives: an $\Omega(\min(\log m, n))$ step lower bound on deterministic and $\Omega(\log m / \log \log m)$ lower bound, for $m \leq n$, on the expected step complexity of randomized implementations. These bounds, however, use a different proof technique, which is specifically tailored for the semantics of the particular objects, and does not seem to generalize to the restricted-use versions of arbitrary perturbable objects. Moreover, they neither prove space-complexity lower bounds nor consider implementations from arbitrary primitives.

## 2 Model and Definitions

A shared-memory system consists of $n$ asynchronous processes $p_1, \ldots, p_n$ communicating by applying primitive operations (primitives) on shared base objects. An application of each such primitive is a shared memory event. A step taken by a process consists of local computation followed by one shared memory event.

A primitive is nontrivial if it may change the value of the base object to which it is applied, e.g., a write or a read-modify-write, and trivial otherwise, e.g., a read. Let $o$ be a base object that is accessed with two primitives $f$ and $f'$; $f$ overwrites $f'$ on $o$ [10], if starting from any value $v$ of $o$, applying $f'$ and then $f$ results in the same value as applying just $f$, using the same input parameters (if any) in both cases. A set of primitives is historyless if all the nontrivial primitives in the set overwrite each other; we also require that each such primitive overwrites itself. A set that includes the write and swap primitives is an example of a historyless set of primitives.
Executions and Operations: An execution fragment is a sequence of shared memory events applied by processes. An execution fragment is $p_i$-free if it contains no steps of process $p_i$. An execution is an execution fragment that starts from an initial configuration (in which all shared variables and processes' local states assume their initial values).

An operation instance of an operation $O_p$ on an implemented object is a subsequence of an execution, in which some process $p_i$ performs the operation $O_p$ on the object. The primitives applied by the operation instance may depend on the values of the shared base objects before this operation instance starts and during its execution ($p_i$'s steps may be interleaved with steps of other processes). Most implementations are required to be linearizable [13]. An implementation is obstruction-free [12] if a process terminates its operation instance if it runs in isolation long enough.

A process $p$ is active after execution $\alpha$ if $p$ is in the middle of performing an operation instance, i.e., $p$ has applied at least one event of the operation instance in $\alpha$, but the instance is not complete in $\alpha$. Let $active(\alpha)$ denote the set of processes that are active after $\alpha$. If $p$ is not active after $\alpha$, we say that $p$ is idle after $\alpha$. A base object $o$ is covered after an execution $\alpha$ if there is a process $p$ in the configuration resulting from $\alpha$ that has a nontrivial event about to access $o$; we say that $p$ covers $o$ after $\alpha$.

Restricted-Use Objects: Our main focus in this paper is on objects that support restricted usage. One example of such objects are objects that have a limit on the number of operation instances that can be performed on them, as captured by the following definition. An $m$-limited-use object is an object that allows at most $m$ operation instances; $m$ is the limit of the object.

Another type of objects with restricted usage are objects that have a value associated with their state which cannot exceed some bound. Examples are bounded max-registers and bounded counters [3], whose definitions appear in the appendix.

We also consider collect and compare-and-swap objects. A collect object provides two operations: a $store(val)$ by process $p_i$ sets $val$ to the latest value for $p_i$. A collect operation $cop$ returns a $view, \langle v_1, \ldots, v_n \rangle$, satisfying the following properties: 1) if $v_j = \bot$, then no $store$ operation by $p_j$ completes before $cop$ starts, and 2) if $v_j \neq \bot$, then $v_j$ is the operand of a $store$ operation $sop$ by $p_j$ that starts before $cop$ completes and there is no store operation by $p_j$ that starts after $sop$ completes and completes before $cop$ starts. A linearizable $b$-valued compare-and-swap object assumes values from $\{1, \ldots, b\}$ and supports the operations $read$ and $CAS(u,v)$, for all $u, v \in \{1, \ldots, b\}$. When the object’s value is $u$, $CAS(u,v)$ changes its value to $v$ and returns $true$; when the object’s value differs from $u$, $CAS(u,v)$ returns $false$ and does not change the object’s value.

3 Lower Bounds for Deterministic Restricted-Use Objects

In this section, we prove lower bounds for obstruction-free implementations of some restricted-use objects. Our starting point is the definition of perturbable objects by Jayanti et al. [15]. Roughly speaking, an object is perturbable if in some class of executions, events applied by an operation of one process influence the response of an operation of another process. The flavor of the argument used by Jayanti et al. to obtain their linear lower bound is that since the perturbed operation needs to return different responses with each perturbation, it must be able to distinguish between perturbed executions, implying that it must perform an increasing number of accesses to base objects.

Following is the formal definition of perturbable objects.

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Assume the object has a

\[ \lambda \]

order to distinguish between executions in the sequence and be able to return different responses. To achieve our lower bounds we only need to show the existence of a special perturbing not allowing any further operation to affect the object, which rules out a perturbation of these use objects. For such objects, some executions already reach the limit or bound of the object, perturbable object to be perturbable, since this requirement is in general not satisfied by restricted-use objects.

As opposed to the definition of a perturbable object, we do not require every execution of an unbounded length, hence they do not apply in general for restricted-use objects. The linear lower bounds \([15]\) on the space and step complexity of obstruction-free implementations on perturbable objects (as defined in Definition 1 above) are obtained by constructing saturations on perturbable objects.

We observe that \(\alpha \gamma \lambda\) in the above definition is an execution, since no process applies more than a single event in \(\lambda\) and \(p_\ell\) applies no events in \(\lambda\).

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To prove lower bounds for restricted-use objects, we define a class of \(L\)-perturbable objects. As opposed to the definition of a perturbable object, we do not require every execution of an \(L\)-perturbable object to be perturbable, since this requirement is in general not satisfied by restricted-use objects. For such objects, some executions already reach the limit or bound of the object, not allowing any further operation to affect the object, which rules out a perturbation of these executions. To achieve our lower bounds we only need to show the existence of a special perturbing sequence of executions rather than attempting to perturb any execution. The longer the sequence, the higher the lower bound, since the perturbed operation will have to access more base objects in order to distinguish between executions in the sequence and be able to return different responses.

**Definition 1** (See Figure 1.) An object \(O\) is perturbable if there is an operation instance \(op_n\) by process \(p_n\), such that for any \(p_n\)-free execution \(\alpha \lambda\) where no process applies more than a single event in \(\lambda\) and for some process \(p_\ell \neq p_n\) that applies no event in \(\lambda\), there is an extension of \(\alpha, \gamma\), consisting of events by \(p_\ell\), such that \(p_n\) returns different responses when performing \(op_n\) by itself after \(\alpha \lambda\) and after \(\alpha \gamma \lambda\).

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**Definition 2** An object \(O\) is \(L\)-perturbable if there exists an operation instance \(op_n\) by \(p_n\) such that an \(L\)-perturbing execution of \(O\) can be constructed as follows: The empty execution is \(0\)-perturbing. Assume the object has a \((k - 1)\)-perturbing execution \(\alpha_{k - 1} \lambda_{k - 1}\), where no process applies more than a single event in \(\lambda_{k - 1}\).

1. If \(|\lambda_{k - 1}| = n - 1\), then we say that \(\alpha_{k - 1} \lambda_{k - 1}\) is saturated, and the execution \(\alpha_k \lambda_k\) with \(\alpha_k = \alpha_{k - 1}\), \(\lambda_k = \lambda_{k - 1}\) is \(k\)-perturbing.

2. Otherwise, if there exists a process \(p_\ell \neq p_n\) that applies no event in \(\lambda_{k - 1}\) and an extension of \(\alpha_{k - 1}\), \(\gamma\), consisting of events by \(p_\ell\), such that \(p_n\) returns different responses when performing \(op_n\) by itself after \(\alpha_{k - 1} \lambda_{k - 1}\) and after \(\alpha_{k - 1} \gamma \lambda_{k - 1}\), then we define a \(k\)-perturbing execution as follows. Let \(\gamma = \gamma' \epsilon \gamma''\), where \(e\) is the first event of \(\gamma\) such that \(op_n\) returns different responses after \(\alpha_{k - 1} \lambda_{k - 1}\) and after \(\alpha_{k - 1} \gamma' \epsilon \lambda_{k - 1}\). Let \(\lambda\) be any permutation of the events in \(\lambda_{k - 1}\) and the event \(\epsilon\), and let \(\lambda', \lambda''\) be any two sequences of events such that \(\lambda\ = \lambda' \lambda''\). The execution \(\alpha_k \lambda_k\) is \(k\)-perturbing, where \(\alpha_k = \alpha_{k - 1} \gamma' \lambda'\) and \(\lambda_k = \lambda''\).

If an object is \(L\)-perturbable, then, starting from the initial configuration, we may construct a sequence of \(L + 1\) perturbing executions, \(\alpha_k \lambda_k\), for \(0 \leq k \leq L\), each of which extending its
predecessor perturbing execution. If for some $i$, $\alpha_i \lambda_i$ is saturated, then we cannot further extend the sequence of perturbing executions since we do not have available processes to perform the perturbation. However, in this case we have lower bounds that are linear in $n$. For presentation simplicity, we assume in this case that the rest of the sequence’s perturbing executions are identical to $\alpha_i \lambda_i$.

Definition 2 allows flexibility in determining which of the events of $\lambda_{k-1}$ are contained in $\lambda_k$ and which are contained in $\alpha_k$. We use this flexibility to prove lower bounds on the step, space and stall-complexity of $L$-perturbable objects.

The definition implies that every perturbable object is $L$-perturbable for every integer $L \geq 0$, hence, the class of $L$-perturbable objects generalizes the class of perturbable objects. On the other hand, there are $L$-perturbable objects that are not perturbable; for example, a $b$-bounded $n$-process max register, for $b \in \text{poly}(n)$, is not perturbable in general, by the algorithm of [3]. That is, the class of perturbable objects is a proper subset of the class of $L$-perturbable objects.

Lemma 5 (in the appendix) establishes that several common restricted-use objects are $L$-perturbable, where $L$ is a function of the limit on the number of different operations that may be applied to them. The challenge in proving this lemma is in increasing $L$, which later translates to higher lower bounds. The specific bounds are summarized in Table 1.

### 3.1 Lower bounds for implementations using historyless objects

We define the concept of an access-perturbation sequence, and prove a step-complexity lower bound for objects that admit such a sequence.

**Definition 3** (See Figure 2.) An access-perturbation sequence of length $L$ of an operation instance $op_n$ by process $p_n$ on an object $O$ is a sequence of executions $\{\alpha_r \lambda_r \phi_r\}_{r=0}^L$, such that $\alpha_0 \lambda_0$ is empty, $\phi_0$ is an execution of $op_n$ by $p_n$ starting from the initial configuration, and for every $r$, $1 \leq r \leq L$, the following properties hold:

1. The execution $\alpha_r \lambda_r$ is $p_n$-free.

2. In $\phi_r$, process $p_n$ runs solo after $\alpha_r \lambda_r$ until it completes the operation instance $op_n$, in the course of which it accesses the base objects $B_1^1, \ldots, B_r^{p_r}$.

3. $\lambda_r$ consists of $j_r \geq 0$ nontrivial events applied by $j_r$ distinct processes, $p_1^1, \ldots, p_{j_r}^1$, to distinct base objects $O_1^1, \ldots, O_{j_r}^1$, respectively, all of which are accessed by $p_n$ in $\phi_r$. If $j_r = n - 1$, we say that $\alpha_r \lambda_r \phi_r$ is saturated.

4. (a) If $\alpha_{r-1} \lambda_{r-1} \phi_{r-1}$ is saturated, then we let $\alpha_r = \alpha_{r-1}$, $\lambda_r = \lambda_{r-1}$ and $\phi_r = \phi_{r-1}$.

   (b) Otherwise, we let $\alpha_r = \alpha_{r-1} \gamma_{r-1}' \lambda_{r-1}'$, and $\lambda_r = \lambda_{r-1}' \epsilon_r$, where $\lambda_{r-1}'$ is the subset of $\lambda_{r-1}$ containing all events to base objects that are not accessed by $p_n$ in $\phi_r$, $\lambda_{r-1}''$ is the subset of $\lambda_{r-1}$ containing all events to base objects that are accessed by $p_n$ in $\phi_r$, and $\gamma_{r-1}' \epsilon_r$ is an execution fragment by a process $p_{r-1}$ not taking steps in $\lambda_{r-1}$, where $\epsilon_r$ is its first nontrivial event to a base object in $\{B_1^{r-1}, \ldots, B_{r-1}^{p_{r-1}}\} \setminus \{O_1^{r-1}, \ldots, O_{r-1}^{p_{r-1}}\}$.

Next, we prove a step lower bound for implementations that have an access-perturbation sequence. If the sequence is saturated, then the lower bound is linear in the number of processes, otherwise it is logarithmic in the length of the sequence. Our goal is to prove that $p_n$ has to access
Figure 2: An access-perturbation sequence of length $L$: the above describes the executions for every $r$, $1 \leq r \leq L$. Notice that $\alpha_r \lambda_r$ is $p_n$-free for every $r$.

a large number of base objects as it runs solo while performing an instance $op_n$ of $Op$ in one of the executions of $op_n$’s access-perturbation sequence. Let $\pi_r$ denote the sequence of base objects accessed by $p_n$ in $\phi_r$, in the order of their first access in $\phi_r$; $\pi_r$ is $p_n$’s solo path in $\phi_r$. If all the objects accessed in $\lambda_{r-1}$ are also in $\lambda_r$, i.e., $p_n$ accesses them also in $\phi_r$, then $\lambda_r = \lambda_{r-1} \epsilon_r$. However, the application of $\epsilon_r$ may have the undesirable effect (from the perspective of an adversary) of making $\pi_r$ shorter than $\pi_{r-1}$: $p_n$ may read the information written by $p_{\epsilon_r}$ and avoid accessing some other objects that were previously in $\pi_{r-1}$.

To overcome this difficulty, we employ the backtracking covering technique [5,11]. The observation underlying this technique is that objects that are in $\pi_{r-1}$ will be absent from $\pi_r$ only if the additional object to which $p_{\epsilon_r}$ applies the nontrivial event $\epsilon_r$ precedes them in $\pi_{r-1}$. Thus the set of objects along $\pi_r$ that are covered after $\alpha_r \lambda_r$ is ‘closer’, in a sense, to the beginning of $p_n$’s solo path in $\phi_{r-1}$. It follows that if there are many access-perturbation sequence executions $r$ for which $|\pi_r| < |\pi_{r-1}|$, then one of the solo paths $\pi_r$ must be ‘long’.

To capture this intuition, we define $\Psi$, a monotonically-increasing progress function of $r$. $\Psi_r$ is a $(\log L)$-digit binary number defined as follows. Bit 0 (the most significant bit) of $\Psi_r$ is 1 if and only if the first object in $\pi_r$ is covered after $\alpha_r$ (by one of the events of $\lambda_r$); bit 1 of $\Psi_r$ is 1 if and only if the second object in $\pi_r$ exists and is covered after $\alpha_r$, and so on. Note that we do not need to consider paths that are longer than $\log L$. If such a path exists, the lower bound clearly holds.

To construct the $r$’th access-perturbation sequence execution, we deploy a free process, $p_{\epsilon_r}$, and let it run solo until it is about to write to an uncovered object, $O$, along $\pi_r$. (Since the sequence is not saturated, it follows from Property 4(b) of Definition 3 that such $p_{\epsilon_r}$ and $O$ exist.) In terms of $\Psi$, this implies that the covering event $\epsilon_r$ might flip some of the digits of $\Psi_{r-1}$ from 1 to 0. But $O$ corresponds to a more significant digit, and this digit is flipped from 0 to 1, hence $\Psi_r > \Psi_{r-1}$ must hold. Thus we can construct executions $\alpha_r \lambda_r \phi_r$, for $1 \leq r \leq L$, in each of which $\Psi_r$ increases. It follows that $\Psi_r = L - 1$ must eventually hold, implying that $\pi_r$’s length is $\Omega(\log L)$.

**Theorem 1** Let $A$ be an $n$-process obstruction-free implementation of an $L$-perturbable object $O$ from historyless primitives. Then $A$ has an execution in which some process accesses $\Omega(\min(\log L, n))$ distinct base objects during a single operation instance.

**Proof:** Lemma 6 (in the appendix) shows that any implementation of $O$ from historyless primitives has an access-perturbation sequence of length $L \geq 1$, $\{\alpha_r \lambda_r \phi_r\}^{L}_{r=0}$. If the sequence is saturated, then Definition 3 immediately implies that $p_n$ accesses $n - 1$ distinct base objects in the course of performing $\phi_r$, and the lower bound holds. Otherwise, we show that $op_n$ accesses $\Omega(\log L)$ distinct base objects in one of these executions.

Let $\pi_r = B^1_r \ldots B^r_{i_r}$ denote the sequence of all distinct base objects accessed by $p_n$ in $\phi_r$ (after $\alpha_r \lambda_r$) according to Property 2 of Definition 3, and let $S_{\pi_r}$ denote the set of these base objects. Let
$S^C_r = \{O^1_r, \ldots, O^b_r\}$ be the set of base objects defined in Property 3 of Definition 3. Observe that, by Property 3, $S^C_r \subseteq S_{\pi_r}$ holds. Without loss of generality, assume that $O^1_r, \ldots, O^b_r$ occur in $\pi_r$ in the order of their superscripts.

In the execution $\alpha_r \lambda_r \phi_r, p_n$ accesses $i_r$ distinct base objects. Thus, it suffices to show that some $i_r$ is in $\Omega(\log L)$. For $j \in \{1, \ldots, i_r\}$, let $b^j_r$ be the indicator variable whose value is 1 if $B^j_r \in S^C_r$ and 0 otherwise. We associate an integral progress parameter, $\Psi_r$, with each $r \geq 0$, defined as follows:

$$\Psi_r = \sum_{j=1}^{i_r} b^j_r \cdot \frac{L}{2^j}.$$  

For simplicity of presentation, and without loss of generality, assume that $L = 2^s$ for some integer $s > 0$, so $s = \log L$. If $i_r > s$ for some $r$ then we are done. Assume otherwise, then $\Psi_r$ can be viewed as a binary number with $s$ digits whose $j$’th most significant bit is 1 if the $j$’th base object in $\pi_r$ exists and is in $S^C_r$, or 0 otherwise. This implies that the number of 1-bits in $\Psi_r$ equals $|S^C_r|$. Our execution is constructed so that $\Psi_r$ is monotonically increasing in $r$ and eventually, for some $r'$, $\Psi_{r'}$ equals $L - 1 = L \sum_{j=1}^{s} \frac{1}{2^j}$. This would imply that $p_n$ accesses exactly $s$ base objects during $\phi_{r'}$ (after $\alpha_r \lambda_r$).

We next show that $\Psi_r > \Psi_{r-1}$, for every $0 < r \leq L$. Since $\alpha_{r-1} \lambda_{r-1} \phi_{r-1}$ is not saturated, by Property 4(b) of Definition 3, there is a process $p_{l_r}$ that takes no steps in $\lambda_{r-1}$, and an execution fragment $\gamma'_e r_r$ of $p_{l_r}$, after $\alpha_{r-1}$, such that $e_r$ is the first nontrivial event of $p_{l_r}$ in $\gamma'_e r_r$ to a base object in $\{B^1_{r-1}, \ldots, B^{i_{r-1}}_{r-1}\} \setminus \{O^1_{r-1}, \ldots, O^{i_{r-1}}_{r-1}\}$. By Property 2 of that definition, this object is accessed by $p_n$ in $\phi_r$. Let $k$ be the index of the object among the objects accessed in $\phi_{r-1}$, i.e., it is $B^k_{r-1}$. This implies that $B^k_{r-1} \in S_{\pi_{r-1}} \setminus S^C_{r-1}$.

As $B^k_{r-1} \notin S^C_{r-1}$, we have $b^k_{r-1} = 0$. Since $e_r$ is the first nontrivial event of $p_{l_r}$ in $\gamma'_e r_r$ to a base object in $S_{\pi_{r-1}} \setminus S^C_{r-1}$, we have that the values of objects $B^1_{r-1} \cdots B^{k-1}_{r-1}$ are the same after $\alpha_{r-1} \lambda_{r-1}$ and $\alpha_r \lambda_r$. It follows that $b^j_{r-1} = b^j_r$ for $j \in \{1, \ldots, k-1\}$. This implies, in turn, that $B^k_{r-1} = B^k_r$.

As $B^k_r \in S^C_r$, we have $b^k_r = 1$. In the appendix, we use the observation that $b^k_{r-1} = 0$ to prove that $\Psi_r > \Psi_{r-1}$ (Lemma 7). Since $\Psi_0 = 0$ and since $\Psi_r$ strictly grows with $r$ and can never exceed $L - 1$, it follows that $\Psi_L = L - 1$, which concludes the proof.

The specific lower bounds appear in Theorem 8 in the appendix, and are summarized in Table 1. To prove space-complexity lower bounds on $L$-perturbable objects, we construct perturbing sequences in which many objects are covered; not all of them are necessarily accessed by the reader, but, nevertheless, they must be distinct, giving a lower bound on the number of base objects. In the appendix, we define a cover-perturbation sequence, which immediately yields a space-complexity lower bound for objects that admit such a sequence, linear in its length (Theorem 9), and prove that every $L$-perturbable object has such a sequence (Lemma 10). Together with Lemma 5, this gives the specific lower bounds stated in Theorem 11 (in the appendix) and summarized in Table 1.

### 3.2 Lower bounds for implementations using arbitrary primitives

The number of steps performed by an operation, as we have measured for implementations using only historyless objects, is not the only factor influencing the performance of an operation. The performance of a concurrent object implementation is also influenced by the extent to which multiple processes simultaneously access widely-shared memory locations. Dwork et al. [8] introduced a formal model to capture such contention, taking into consideration both the number of steps taken
by a process and the number of stalls it incurs as a result of memory contention with other processes. More formally, an event \( e \) applied by a process \( p \) to object \( O \) in an execution \( \alpha \) incurs \( k \) memory stalls if it is immediately preceded by \( k \) events by distinct processes different than \( p \) that apply nontrivial primitives to \( O \).

Our next result shows a lower bound on implementations using arbitrary read-modify-write primitives. Its proof employs a variation of the backtracking covering technique, similar to the proof of Theorem 1. The earlier proof uses access-perturbable sequence of executions, in which each new execution deploys a process to cover an object that is not covered in the preceding execution. Such a series of executions cannot, in general, be constructed for algorithms that may use arbitrary primitives. Instead, the proof constructs a series of executions in which each new execution deploys a process that covers some object along \( p_n \)'s path.

**Definition 4** An access-stall perturbation sequence of length \( L \) of an operation instance \( \text{op}_n \) by process \( p_n \) on an object \( O \) is a sequence of executions \( \alpha, \sigma_{r,1} \cdots \sigma_{r,j_r}, \rho_r \), such that \( \alpha_0 \) is empty, \( j_0 = 0 \), \( \rho_0 \) is an execution of \( \text{op}_n \) by \( p_n \) starting from the initial configuration, and for every \( r \), \( 1 \leq r \leq L \), the following properties hold:

1. \( \alpha_r \) is \( p_n \)-free,
2. in \( \rho_r \), process \( p_n \) runs solo until it completes the operation instance \( \text{op}_n \); in this instance, \( p_n \) accesses the base objects \( B^1_r, \ldots, B^r_r \),
3. there is a subsequence \( O^1_r, \ldots, O^{j_r}_r \) of disjoint objects in \( B^1_r, \ldots, B^r_r \) and disjoint nonempty sets of processes \( S^1_r, \ldots, S^{j_r}_r \) such that, for \( j = 1, \ldots, j_r \),
   - each process in \( S^j_r \) covers \( O^j_r \) after \( \alpha_r \), and
   - in \( \sigma_{r,j} \), process \( p_n \) applies events until it is about to access \( O^j_r \) for the first time, then each of the processes in \( S^j_r \) accesses \( O^j_r \), and, finally, \( p_n \) accesses \( O^j_r \).
4. let \( \lambda_{r-1} \) be the subsequence of events by the processes in \( S^1_{r-1} \cup \cdots \cup S^{j_{r-1}}_{r-1} \) that are applied in \( \sigma_{r-1,1} \cdots \sigma_{r-1,j_{r-1}} \), then \( \alpha_{r-1} \lambda_{r-1} \) is an \( r-1 \)-perturbing execution; if \( \alpha_{r-1} \lambda_{r-1} \) is saturated, then we say that \( \alpha_{r-1} \sigma_{r-1,1} \cdots \sigma_{r-1,j_{r-1}} \rho_{r-1} \) is saturated,
5. If \( \alpha_{r-1} \sigma_{r-1,1} \cdots \sigma_{r-1,j_{r-1}} \rho_{r-1} \) is saturated, then the \( r \)’th execution in the access-stall perturbation sequence is defined as identical to it. Otherwise, the following holds: \( O^{i_r}_r = B^k_{r-1} \), for some \( 1 \leq k \leq i_{r-1} \); \( B^i_{r-1} = B^i_{r-1} \), for all \( i \in \{1, \ldots, k\} \); \( O^1_{r-1} = O^i_{r-1} \) and \( S^1_{r-1} = S^i_{r-1} \) for all objects \( O^1_{r-1} \) that precede \( B^k_{r-1} \) in the sequence \( B^1_{r-1}, \ldots, B^k_{r-1} \); and either \( B^k_{r-1} \not\in \{O^1_{r-1}, \ldots, O^{i_{r-1}}_{r-1}\} \) or \( O^{j_{r-1}}_{r-1} = O^{j_{r-1}}_{r-1} \) and \( |S^{j_{r-1}}_{r-1}| = |S^{j_{r-1}}_{r-1}| + 1 \).

**Theorem 2** Let \( A \) be an \( n \)-process obstruction-free implementation of an \( L \)-perturbable object \( O \) from any read-modify-write primitives. Then \( A \) has an execution in which some process either accesses \( \Omega(\min(\log L, n)) \) distinct base objects or incurs \( \Omega(\min(\log L, n)) \) memory stalls, during a single operation instance.

**Proof:** For simplicity and without loss of generality, assume that \( L = 2^{2s} \) for some integer \( s \). If \( A \) has an execution in which some process accesses \( s \) distinct base objects during a single operation instance, then the theorem holds. Otherwise, it can be shown that \( A \) has an access-stall perturbation
sequence of length $L$ (Lemma 12 in the appendix). If one of these executions, $\alpha_r \sigma_{r,1} \cdots \sigma_{r,j}, \rho_r$, for some $r \leq L$, is saturated, then it follows from Definition 4 that $p_n$ incurs $n - 1$ memory stalls in the course of $\sigma_{r,1} \cdots \sigma_{r,j}$, and the theorem holds. We therefore assume in the following that none of the executions in $A$’s access-stall perturbation sequence is saturated. We will prove that $p_n$ incurs $\Omega(s)$ memory stalls in one of these executions.

For $i \in \{1, \ldots, i_r\}$, let variable $n^i_r$ be defined as follows:

$$n^i_r = \begin{cases} |S^m_r|, & \text{if } \exists m \in \{1, \ldots, j_r\} : B^i_r = O^m_r, \\ 0, & \text{otherwise.} \end{cases}$$

(1)

Let $N_r = \sum_{i=1}^{i_r} n^i_r$. Thus, it suffices for the proof to show that one of these executions has $N_r = \Omega(s)$. We associate the following integral progress parameter, $\Phi_r$, with each execution $r \geq 0$:

$$\Phi_r = \sum_{i=1}^{i_r} n^i_r \cdot s^{s-i}. \tag{2}$$

If $n^i_r \geq s - 1$ for some $0 \leq r \leq L$ and $i \in \{1, \ldots, i_r\}$, then we are done, since clearly $N_r \geq s - 1$ holds in this case. Assume otherwise, then $\Phi_r$ can be viewed as an $s$-digit number in base $s$ whose $i$th most significant digit is $0$ if $i > i_r$ or equals the number of processes in $S^1_r, \ldots, S^3_r$ covering $B^i_r$ after $\alpha_r$ otherwise.

From the last property of Definition 4, $O^{i_r+1}_r = B^k_r$, for some $1 \leq k \leq i_r$ and, moreover, $B^k_{i+1} = B^i_r$ for $i \in \{1, \ldots, k\}$, $n^i_{i+1} = n^i_r$, for $i \in \{1, \ldots, k-1\}$, $n^k_{i+1} = n^k_r + 1$, and $n^{i+1}_{i+1} = 0$ for $i \in \{k + 1, \ldots, i_r\}$. Based on this, we prove that $\Phi_{i+1} > \Phi_r$ (Lemma 13 in the appendix).

Since the sequence $\Phi_1, \ldots, \Phi_L$ is strictly growing, each $\Phi_r$ is unique. By the definition of $\Phi$, each value $\Phi_r$ corresponds to a different partitioning of integer $N_r$ to the values of the $s$ digits of $\Phi_r$. What is the maximum number $N$ of different executions $r$ for which $N_r \leq s$ holds? $N$ is at most the number of distinguishable partitions of up to $s$ identical balls into $s$ bins. Let $A_{b,c}$ be the number of distinguishable partitions of $b$ identical balls into $c$ bins, then:

$$N \leq \sum_{j=0}^{s} A_{j,s} = A_{s,s+1} = \binom{2s}{s} = \left( \frac{\log L}{\log L/2} \right) = \Theta \left( \frac{4\log L/2}{\sqrt{\pi \log L/2}} \right) = \Theta \left( \frac{L}{\sqrt{\pi \log L/2}} \right) < L.$$

Where the one-before-last equality above follows from Stirling’s approximation and the error of the approximation ratio \(\frac{\log L}{\log L/2}/\frac{4\log L/2}{\sqrt{\pi \log L/2}}\) is inversely proportional to $s$ [9, page 75]. Thus, for all $L \geq 4$, there is an execution $\alpha_{r'} \sigma_{r',1} \cdots \sigma_{r',j'}, \rho_{r'}$ such that $N_{r'} > s$ holds.

Together with Lemma 5, we obtain the specific bounds presented in Theorem 14 (in the appendix), and summarized in Table 1.

4 Lower Bound for Randomized Approximate Counters

Proving lower bounds for randomized implementations of concurrent objects is more difficult, due to the extra flexibility these implementations have. We were not able to prove general lower bounds for a class of objects, but we take a first step in this direction by proving a lower bound for a specific, but very useful, object, namely an approximate counter. This object allows some error
in the operations applied to them. We consider two variants, depending on whether the error is additive or multiplicative.

We assume an oblivious adversary, which fixes the sequence of process steps in advance, without being able to predict the coin-flips of the processes or the progress of the execution; in fact, our adversary does not even require knowledge of the implementation, allowing us to prove the lower bound using Yao’s Principle [16]. We consider deterministic algorithms, since a randomized algorithm can be seen as a weighted average of deterministic ones. A distribution over schedules that gives a high cost on average for any fixed deterministic algorithm, also gives a high cost on average for any randomized algorithm, which also implies that there exists some specific schedule that does so. We will describe an (oblivious) adversary strategy achieving the next lower bound:

**Theorem 3** For any randomized implementation of an m-valued c-multiplicative-accurate counter using historyless primitives for n ≥ m processes, and any fixed ε > 0, there is an oblivious adversary strategy that yields, with probability at least 1 − ε, an execution consisting of at most m − 1 concurrent CounterIncrement operation instances, some of which may be incomplete, followed by a CounterRead operation instance, in which one of the following conditions holds: (a) a constant fraction of the CounterIncrement instances take more than w operations; (b) the value returned by the CounterRead operation instance is not consistent with any linearization of the completed operation instances; or (c) the CounterRead operation instance takes $\Omega\left(\log\log m - \log\log c / \log w\right)$ operations.

We first consider a schedule constructed as follows. Process $p_1$ carries out an operation $\beta_1$ for at most w steps. With probability $p$ for each step, $p_1$ is stopped early and is suspended before it can carry the step out; if the step is not a read operation, this means that the target register is now covered by a pending operation that can be delivered later to overwrite any subsequent work by other processes. Whether $p_1$ completes its operation or not, process $p_2$ is next scheduled to carry out at most w steps, each of which causes $p_2$ to be suspended with probability $p$ as before, and this process is repeated for the remaining processes up through $p_{n-1}$. In this way we assemble a schedule $\Gamma = \beta_1'\beta_2'...\beta_{n-1}'$, where each $\beta_i'$ is an initial prefix of some high-level operation $\beta_i$.

From this schedule we construct a family of schedules $\{\Xi_k\}_{k\geq 0}$, where each $\Xi_k$ consists of an initial prefix of $\Gamma$ of length $k$ (i.e., consisting of $k$ steps), followed by the delivery of all delayed operations from $\Gamma$, and in turn followed by the first $r$ steps of a single operation $\alpha$ executed by $p_n$. Thus each $\Xi_k$ is of the form $\beta_1'\beta_2'...\beta_m'\delta_m\delta_{m-1}...\delta_1\alpha$, where $\delta_i$ is either the delayed operation of $i$ or the empty sequence if there is no such operation, $\alpha$ is the single operation of $p_n$, and $\beta_m'$ is a prefix of $\beta_m$ that makes the initial segment have the correct length.

The proof is based first on bounding the number of distinct values returned by the reader across all the schedules $\Xi_k$ as a function of $p$ and $r$, and then showing that we can select a subset of these schedules that must either violate the restriction to short increment and read operations or return significantly more distinct values. This implies that choosing one of these schedules uniformly at random is likely to hit one of the bad outcomes. The next lemma, whose proof appears in the appendix, bounds the number of distinct return values:

**Lemma 4** Among the schedules $\Xi_k$ above, $\alpha$ returns at most $(1 + 1/p)^r$ distinct values on average, where the average is taken over the random choices of the adversary for when to delay operations.

The key idea is that because $\alpha$ is deterministic, the value it returns can depend only on the values of the at most $r$ registers it reads, and that each register will get at most $1 + 1/p$ values on
average in all the $\Xi_k$ before it becomes covered by some $\delta_i$. This is essentially the same idea as used in [3] for max registers, except that we provide a more careful analysis of the dependence between the number of values found in each registers, because the union bound used in [3] reduces the lower bound by a $\Theta(\log \log m)$ factor that in our case would eliminate the lower bound completely.

Lemma 4 holds for arbitrary sequences of operations. To prove Theorem 3, we show that for the specific case where $p = 1/4w$ and each $\beta_i$ is a CounterIncrement and $\alpha$ is a CounterRead for a $c$-multiplicative-accurate counter, we can pick out a subfamily of executions $\Xi_{k_0}, \Xi_{k_1}, \ldots, \Xi_{k_{\ell-1}}$, where $\ell = \left\lceil \frac{1}{2} \log_2 \sqrt{m} \right\rceil - 1 = \Theta(\log m / \log c)$, such that, on average, a constant fraction of the executions $\Xi_{k_i}$ satisfies one of the conditions in Theorem 3. The full proof appears in the appendix.

If we choose $w$ to match the lower bound on CounterRead, we get a lower bound on the worst-case cost of any $c$-multiplicative-accurate counter operation for fixed $c$ of $\Omega \left( \frac{\log \log m}{\log \log \log m} \right)$. This is much smaller than Jayanti’s lower bound of $\Omega(\log n)$ on randomized $n$-bounded counters [14], which also allows much stronger primitives in the implementation. But the smaller bound is not surprising if one considers that a $c$-multiplicative-accurate counter effectively provides only $\Theta(\log \log m)$ bits of information about the number of increments, compared with $\Theta(\log m)$ for standard counter.

5 Summary

This paper presents lower bounds for concurrent obstruction-free implementations of objects that are used in a restricted manner. (See Table 1 in the introduction.) The step lower-bound on max registers is tight [3] and the step lower bound on randomized counters is almost tight, as there is an $O(\log \log m)$ upper bound [7], under the same adversary model. It is unclear whether the other lower bounds are tight. Another interesting research direction is to devise generic implementations for $L$-perturbable objects. This is of particular interest in the case of randomized implementations, where there is also an important issue of the type of adversary tolerated.

References


A Examples of $L$-Perturbable Objects

A **counter** is a linearizable object that supports a `CounterIncrement` operation and a `CounterRead` operation, which returns the number of `CounterIncrement` operation instances linearized before it. In a $k$-additive-accurate counter, any `CounterRead` operation returns a value within $\pm k$ of the number of `CounterIncrement` operation instances linearized before it. A $c$-multiplicative-accurate counter is a counter for which any `CounterRead` operation returns a value $x$ with $v/c \leq x \leq vc$, where $v$ is the number of `CounterIncrement` operation instances linearized before it.

A **max-register** is a linearizable object that supports a `Write`($v$) operation, which writes the value $v$ to the object, and a `ReadMax` operation, which returns the maximum value written by any `Write` operation instance linearized before it. In the bounded version of these objects, the object is only required to satisfy its specification if its associated value does not exceed a certain threshold. A $b$-bounded max register takes values in $\{0, \ldots, b-1\}$. A $b$-bounded counter is a counter that takes values in $\{0, \ldots, b-1\}$. For a $b$-bounded object $O$, $b$ is the **bound** of $O$.

**Lemma 5**

1. An obstruction-free implementation of a $b$-bounded-value max register is $(b-1)$-perturbable.
2. An obstruction-free implementation of an $m$-limited-use max register is $(m-1)$-perturbable.
3. An obstruction-free implementation of an \( m \)-limited-use counter is \((\sqrt{m} - 1)\)-perturbable.

4. An obstruction-free implementation of a \( k \)-additive-accurate \( m \)-limited-use counter is \((\sqrt{m/k} - 1)\)-perturbable.

5. An obstruction-free implementation of an \( m \)-limited-use \( b \)-valued compare-and-swap object is \((\sqrt{m} - 1)\)-perturbable (if \( b \geq n \)).

6. An obstruction-free implementation of an \( m \)-limited-use collect object is \((m - 1)\)-perturbable.

Proof:

1. Let \( O \) be a \( b \)-bounded-value max register and consider an obstruction-free implementation of \( O \). We show that \( O \) is \((b - 1)\)-perturbable for a \texttt{ReadMax} operation instance \( op_n \) of \( p_n \), by induction, where the base case for \( r = 0 \) is immediate for all objects. We perturb the executions by writing values that increase by one to the max register. This guarantees that \( op_n \) has to return different values each time, while getting closer to the limit of the object as slowly as possible.

Formally, let \( r < b \) and let \( \alpha_{r-1}\lambda_{r-1} \) be an \((r - 1)\)-perturbing execution of \( O \). If \( \alpha_{r-1}\lambda_{r-1} \) is saturated, then, by case (1) of Definition 2, it is also an \( r \)-perturbing execution.

Otherwise, our induction hypothesis is that \( op_n \) returns \( r - 1 \) when run after \( \alpha_{r-1}\lambda_{r-1} \). For the induction step, we build an \( r \)-perturbing execution after which the value returned by \( op_n \) is \( r \). Since \( \alpha_{r-1}\lambda_{r-1} \) is not saturated, there is a process \( p_t \neq p_n \) that does not take steps in \( \lambda_{r-1} \). Let \( \gamma \) be the execution fragment by \( p_t \) where it first finishes any incomplete operation in \( \alpha \) and then performs a \texttt{Write} operation to the max register with the value \( r \leq b - 1 \). Then \( op_n \) returns the value \( r \) when run after \( \alpha_{r-1}\gamma\lambda_{r-1} \), and \( r - 1 \) when run after the \((r - 1)\)-perturbing execution \( \alpha_{r-1}\lambda_{r-1} \). It follows that \( r \)-perturbing executions may be constructed from \( \alpha_{r-1}\lambda_{r-1} \) and \( \gamma \) as specified by Definition 2.

2. The proof for an \( m \)-limited-use max register is the same as that for a \( b \)-bounded value max register. We could even allow writing any increasing sequence of values to the max register rather than only increasing by one, since the limit of the object applies to the number of operations rather than to its value.

3. When \( O \) is an \( m \)-limited-use counter, we use a proof similar to the one we used for a limited-use max register, where we perturb a \texttt{CounterRead} operation \( op_n \) by applying \texttt{CounterIncrement} operations. The subtlety in the case of a counter comes from the fact that a single perturbing operation may not be sufficient for guaranteeing that \( op_n \) returns a different value after \( \alpha_{r-1}\lambda_{r-1} \) and after \( \alpha_{r-1}\gamma\lambda_{r-1} \), since we do not know how many of the \texttt{CounterIncrement} operations by processes that are active after \( \alpha_{r-1} \) were linearized. As there are at most \( r - 1 \) such operations, in order to ensure that different values are returned by \( p_n \) after these two executions, we construct \( \gamma \) by letting the process \( p_t \) apply \( r \) \texttt{CounterIncrement} operations after finishing any incomplete operation in \( \alpha_{r-1} \). This can be done as long as \( r \leq \sqrt{m} \) in order not to pass the limit on the number of operations allowed, which will be \( 1 + \sum_{r=1}^{\sqrt{m}} r = 1 + \frac{\sqrt{m} - 1}{2} \sqrt{m} \leq m \).
4. For a \( k \)-additive-accurate \( m \)-limited-use counter the proof is similar to that of a counter, except that \( p_n \) needs to perform an even larger number of CounterIncrement operations in \( \gamma \), because of the inaccuracy allowed in the returned value of the CounterRead operation \( op_n \). Denote by \( I_r \) the number of CounterIncrement operation instances performed by the perturbing process in iteration \( r \). We have that \( I_1 = k + 1 \) in order for \( op_n \) to return at least 1. We claim that for \( r > 1 \), \( I_r = 2k + r \), and prove this by induction. The operation \( op_n \) run after \( \alpha_{r-1}\lambda_{r-1} \) can return a value which is as large as \( \sum_{j=1}^{r-1} I_j + k \). Therefore, we need the number of complete CounterIncrement operation instances after \( \alpha_{r-1}\gamma\lambda_{r-1} \) to be at least \( \sum_{j=1}^{r-1} I_j + k + (k + 1) \), for \( op_n \) to return at least \( \sum_{j=1}^{r-1} I_j + k + 1 \). Besides the CounterIncrement operation instances in \( \gamma \), at least \( \sum_{j=1}^{r-1} I_j - (r - 1) \) CounterIncrement operation instances have finished, therefore setting \( I_r = 2k + r \) implies that \( op_n \) returns at least \( \sum_{j=1}^{r-1} I_j - (r - 1) + I_j - k \), which is \( \sum_{j=1}^{r-1} I_j + k + 1 \) as needed. 

This claim implies that a \( k \)-additive-accurate \( m \)-limited-use counter is \( (\sqrt[4]{m} - 1) \)-perturbable, because the total number of operation instances will be 

\[
1 + \left( (k + 1) + (2k + 2) + \ldots + \left( 2k + \sqrt[4]{m} - 1 \right) \right) \leq 1 + 2k \left( \sqrt[4]{m} - 1 \right) + \frac{\left( \sqrt[4]{m} - 1 \right) \left( \sqrt[4]{m} - 1 + 1 \right)}{2} \\
\leq 1 + (2\sqrt{m} - 2k) + \frac{m}{2} \\
\leq m,
\]

where the last inequality holds for a large enough \( m \) (\( m \geq 16 \)).

5. Let \( O \) be an \( m \)-limited-use \( b \)-bounded compare-and-swap object, \( b \geq n \). We show that it is \( (\sqrt[4]{m} - 1) \)-perturbable for a read operation instance by \( p_n \), by induction, where the base case for \( r = 0 \) is immediate for all objects. In our construction, all processes except for \( p_n \) perform only CAS operation instances.

Let \( r < \sqrt[4]{m} - 1 \) and let \( \alpha_{r-1}\lambda_{r-1} \) be an \( (r - 1) \)-perturbing execution of \( O \). If \( \alpha_{r-1}\lambda_{r-1} \) is saturated, then, by case (1) of Definition 2, it is also an \( r \)-perturbing execution.

Otherwise, our induction hypothesis is that \( \alpha_{r-1}\lambda_{r-1} \) includes at most \( \sum_{i=1}^{r-1} i^2 \) CAS operation instances. Let \( u \) be the value returned by \( op_n \) after \( \alpha_{r-1}\lambda_{r-1} \), and let \( j \) denote the number of processes that apply events in \( \lambda_{r-1} \) and let \( p_{r-1}^1, \ldots, p_{r-1}^j \) be these processes. Let \( \xi \) be an execution fragment that follows \( \alpha_{r-1} \) in which all active processes other than \( p_{r-1}^1, \ldots, p_{r-1}^j \) finish any incomplete operation instances they started in \( \alpha_{r-1} \). For \( k \in \{1, \ldots, j\} \), let \( (u_k, v_k) \) denote the operands of the last CAS operation instance started by \( p_{r-1}^k \) in \( \alpha_{r-1} \). Since \( \alpha_{r-1}\lambda_{r-1} \) is not saturated and since \( r - 1 < \sqrt[4]{m} - 1 \), there is a process \( p_\ell \notin \{p_{r-1}^1, \ldots, p_{r-1}^j\} \cup \{p_n\} \) and, moreover, there is a value \( v \in \{1, \ldots, n\} \setminus \{u, u_1, \ldots, u_j\} \).

Denote by \( \beta \) the sequence of operation instances \( CAS(u, v)CAS(v_1, v) \ldots CAS(v_j, v) \), denote by \( \beta^r \) the sequence of operation instances resulting from concatenating \( r \) copies of \( \beta \) and let \( \gamma = \xi \beta^r \).

We claim that \( O \)'s value after \( \alpha_{r-1}\gamma \lambda_{r-1} \) is \( v \). Consider \( O \)'s value after \( p_\ell \) executes \( \beta \) once after \( \alpha_{r-1}\xi \). There are two possibilities: either \( O \)'s value is \( v \) (in which case it remains \( v \) also after \( \alpha \gamma \)), or, otherwise, all the CAS instances in \( \beta \) failed, implying that one or more of the operation instances performed by \( p_{r-1}^1, \ldots, p_{r-1}^j \) are linearized during the execution of \( \gamma \). In
the latter case, consider \( O \)'s value after \( p_\ell \) executes \( \beta \) twice after \( \alpha_{r-1} \xi \). Once again, either \( O \)'s value is \( v \) (and remains \( v \) also after \( \alpha_{r-1} \gamma \)), or, otherwise, additional operation instances performed by \( p^1_{r-1}, \ldots, p^j_{r-1} \) are linearized during the second execution of \( \beta \). Applying this argument iteratively and noting that \( j \leq r - 1 \), by construction, establishes our claim.

Consider the execution \( \alpha_{r-1} \gamma \lambda_{r-1} \phi \), where \( \phi \) is an execution of \texttt{read} by \( p_n \). Then \( \phi \) must return \( v \), whereas an execution of \texttt{read} by \( p_n \) after \( \alpha_{r-1} \lambda_{r-1} \) returns \( u \neq v \). Execution \( \alpha_{r} \lambda_{r} \) can now be constructed as in the proofs for limited use max registers and counters. The number of operation instances applied by \( p_\ell \) in \( \gamma \) is \((j + 1) \cdot r \leq r^2 \). Since \( O \) allows only \( m \) operation instances, this implies that the sequence can have length \( \sqrt{m} - 1 \), because the total number of operation instances will be \( 1 + \sum_{i=1}^{\sqrt{m} - 1} r^2 < m \).

6. Let \( O \) be an \( m \)-limited-use collect object and consider an obstruction-free implementation of \( O \). We show that \( O \) is \((m - 1)\)-perturbable for a \texttt{collect} operation instance \( op_n \) of \( p_n \), by induction, where the base case for \( r = 0 \) is immediate for all objects. We perturb the executions by having processes store values that change their collect component. This guarantees that \( op_n \) has to return different values each time, while getting closer to the limit of the object as slowly as possible.

Formally, let \( r < m \) and let \( \alpha_{r-1} \lambda_{r-1} \) be an \((r - 1)\)-perturbing execution of \( O \). If \( \alpha_{r-1} \lambda_{r-1} \) is saturated, then, by case (1) of Definition 2, it is also an \( r \)-perturbing execution.

Otherwise, Let \( V = \langle v_1, \ldots, v_n \rangle \) denote the value that is returned by a \texttt{collect} operation by \( p_n \) after \( \alpha_{r-1} \lambda_{r-1} \). Since \( \alpha_{r-1} \lambda_{r-1} \) is not saturated, there is a process \( p_\ell \neq p_n \) that does not take steps in \( \lambda_{r-1} \). Let \( \gamma \) be the execution fragment by \( p_\ell \) where it first finishes any incomplete operation in \( \alpha \) and then applies an \texttt{update}(\( v'_\ell \)) operation operation to \( O \), for some \( v'_\ell \neq v_\ell \). Then \( op_n \) must return different values when run after \( \alpha_{r-1} \gamma \lambda_{r-1} \) and after the \((r - 1)\)-perturbing execution \( \alpha_{r-1} \lambda_{r-1} \). It follows that \( r \)-perturbing executions may be constructed from \( \alpha_{r-1} \lambda_{r-1} \) and \( \gamma \) as specified by Definition 2.

\[ \square \]

**B Proofs Omitted from Section 3**

**Lemma 6** An \( L \)-perturbable object implementation from historyless primitives has an access-perturbation sequence of length \( L \).

**Proof:** Let \( O \) be an \( L \)-perturbable object implementation from historyless primitives. We show that it has an access-perturbation sequence of length \( L \), for the operation \( op_n \) as defined in Definition 3. The proof is by induction, where we prove the existence of the execution \( \alpha_{r} \lambda_{r} \phi_{r} \), for every \( r, 0 \leq r \leq L \). To allow the proof to go through, in addition to proving that the execution \( \alpha_{r} \lambda_{r} \phi_{r} \) satisfies the 4 conditions of Definition 3, we will prove that \( \alpha_{r} \lambda_{r} \) is \( r \)-perturbing.

For the base case, \( r = 0 \), \( \alpha_{0} \lambda_{0} \) is empty and \( \phi_{0} \) is an execution of \( op_n \) starting from the initial configuration. Moreover, the empty execution is \( 0 \)-perturbing. We next assume the construction of the sequence up to \( r - 1 \leq L \) and construct the next execution \( \alpha_{r} \lambda_{r} \phi_{r} \) as follows.

By the induction hypothesis, the execution \( \alpha_{r-1} \lambda_{r-1} \) is \((r - 1)\)-perturbing. If \( \alpha_{r-1} \lambda_{r-1} \) is saturated, then, by case (1) of Definition 2, \( \alpha_{r} = \alpha_{r-1} \), \( \lambda_{r} = \lambda_{r-1} \) and \( \alpha_{r} \lambda_{r} \) is \( r \)-perturbing.
Moreover, by property 4(a) of Definition 3, \( \alpha_r \lambda_r \phi_r \) is the \( r \)’th access-perturbation execution, where \( \phi_r = \phi_{r-1} \).

Assume otherwise. Then, by property 4(b) of Definition 3, there is a process \( p_{t_r} \neq p_n \) that does not take steps in \( \lambda_{r-1} \), for which there is an extension of \( \alpha_{r-1} \), \( \gamma_r \), consisting of events by \( p_{t_r} \), such that \( p_n \) returns different responses when performing \( op_n \) by itself after \( \alpha_{r-1} \lambda_{r-1} \) and after \( \alpha_{r-1} \gamma_r \lambda_{r-1} \). As per Definition 2, let \( \gamma_r = \gamma'_r e_r \gamma''_r \), where \( e_r \) is the first event of \( \gamma \) such that \( op_n \) returns different responses after \( \alpha_{r-1} \lambda_{r-1} \) and after \( \alpha_{r-1} \gamma'_r e_r \lambda_{r-1} \). Clearly \( e_r \) is a nontrivial event.

Denote by \( \phi_r \) the execution of \( op_n \) by \( p_n \) after \( \alpha_{r-1} \gamma'_r e_r \lambda_{r-1} \). Since \( op_n \) returns different values after \( \alpha_{r-1} \lambda_{r-1} \) and after \( \alpha_{r-1} \gamma'_r e_r \lambda_{r-1} \), and since the implementation uses only historyless primitives, this implies that \( e_r \) is applied to some base object \( B \) not in \( \{ \lambda_{r-1}^{1}, \ldots, \lambda_{r-1}^{j_r-1} \} \) that is accessed by \( p_n \) in \( \phi_r \).

We define \( \lambda'_{r-1} \) to be the subsequence of \( \lambda_{r-1} \) containing all events to base objects that are not accessed by \( p_n \) in \( \phi_r \), and \( \lambda''_{r-1} \) to be the subsequence of \( \lambda_{r-1} \) containing all events to base objects that are accessed by \( p_n \) in \( \phi_r \). We then define \( \alpha_r = \alpha_{r-1} \gamma'_r \lambda'_{r-1} \), \( \lambda_r = \lambda''_{r-1} e_r \) and show that \( \alpha_r \lambda_r \phi_r \) satisfies the properties of Definition 3.

We first observe that \( \alpha_r \lambda_r \phi_r \) is a well defined execution, since the execution fragment \( \gamma'_r \) by \( p_{t_r} \) is performed after \( \alpha_{r-1} \lambda_{r-1} \), and all operations in \( \lambda_{r-1} \) are nontrivial events to distinct base objects none of which is by \( p_{t_r} \). It follows that \( \alpha_r \lambda_r \) and \( \alpha_{r-1} \gamma'_r e_r \lambda_{r-1} \) are indistinguishable to \( p_n \), hence \( \phi_r \) is a solo execution of \( op_n \) by \( p_n \) after both executions.

Property 1 holds since \( \alpha_r \lambda_r \) is \( p_n \)-free by construction, and \( \phi_r \) is a solo execution fragment by \( p_n \) in which it performs \( op_n \), so Property 2 holds. To show Property 3, we observe that \( \alpha_r \lambda_r \) is indistinguishable to \( p_n \) from \( \alpha_{r-1} \gamma'_r e_r \lambda_{r-1} \) and hence \( p_n \) accesses the base object \( B \) in \( \phi_r \). Finally, Property 4 follows by construction.

We conclude the proof by claiming that \( \alpha_r \lambda_r \) is \( r \)-perturbing, which follows from its construction and Definition 2.

**Lemma 7** \( \Psi_r > \Psi_{r-1} \).

**Proof:**
\[
\Psi_r = \sum_{j=1}^{i_r} b^j \cdot \frac{L}{2^j} = \sum_{j=1}^{k-1} b^j \cdot \frac{L}{2^j} + b^k \cdot \frac{L}{2^k} + \sum_{j=k+1}^{i_r} b^j \cdot \frac{L}{2^j} = \sum_{j=1}^{k-1} b^j \cdot \frac{L}{2^j} + \frac{L}{2^k} + \sum_{j=k+1}^{i_r} b^j \cdot \frac{L}{2^j} \geq \sum_{j=1}^{k-1} b^j \cdot \frac{L}{2^j} + \frac{L}{2^k} \sum_{j=k+1}^{i_r} b^j \cdot \frac{L}{2^j} = \Psi_{r-1},
\]
where the last equality is based on the observation that \( b^k \) is 0.

**Theorem 8** An \( n \)-process obstruction-free implementation of an \( m \)-limited-use max register, \( m \)-limited-use counter, \( m \)-limited-use \( b \)-valued compare-and-swap object or an \( m \)-limited-use collect object from historyless primitives has an operation instance requiring \( \Omega(\min(\log m, n)) \) steps. An obstruction-free implementation of a \( b \)-bounded max register from historyless primitives has an operation instance requiring \( \Omega(\min(\log b, n)) \) steps. An obstruction-free implementation of a \( k \)-additive-accurate \( m \)-limited-use counter from historyless primitives has an operation instance requiring \( \Omega(\min(\log m - \log k, n)) \) steps.
Definition 5 A cover-perturbation sequence of length \(1 \leq L \leq n-1\) of an operation instance \(op_n\) by process \(p_n\) on an object \(O\) is a sequence of executions \(\{\alpha_r, \lambda_r, \phi_r\}_{r=0}^{L}\), such that \(\alpha_0\lambda_0\) is empty, \(\phi_0\) is an execution of \(op_n\) by \(p_n\), and for every \(r, 1 \leq r \leq L\), the following hold.

1. The execution \(\alpha_r\lambda_r\) is \(p_n\)-free.
2. In \(\phi_r\), process \(p_n\) runs solo after \(\alpha_r\lambda_r\) until it completes \(op_n\).
3. In \(\lambda_r\), distinct processes \(q_1, \ldots, q_r\) each apply a nontrivial event to distinct base objects \(O_1, \ldots, O_r\), respectively.
4. \(|active(\alpha_r\lambda_r)| \leq r\).

The space lower bound for cover-perturbable objects follows immediately from Property 3:

Theorem 9 Let \(A\) be an \(n\)-process obstruction-free implementation of an object \(O\) from historyless primitives. If \(A\) has a cover-perturbation sequence of length \(L\), then \(A\) has an execution in which \(L\) distinct base objects are accessed.

The next lemma shows that every \(L\)-perturbable object has a cover-perturbation sequence of length \(L\).

Lemma 10 An \(L\)-perturbable object has a cover-perturbation sequence of length \(L\).

Proof: We show that the object has a cover-perturbation sequence of length \(L\) for operation instance \(op_n\) as defined in Definition 5. The proof is by induction, where we prove the existence of the execution \(\alpha_r\lambda_r\phi_r\), for every \(r, 0 \leq r \leq L\). To allow the proof to go through, in addition to proving that the execution \(\alpha_r\lambda_r\phi_r\) satisfies the four conditions of Definition 5, we will prove that \(\alpha_r\lambda_r\) is \(r\)-perturbing.

For the base case, \(r = 0\), \(\alpha_0\lambda_0\) is empty and \(\phi_0\) is an execution of \(op_n\) starting from the initial configuration. Moreover, the empty execution is 0-perturbing. We next assume the construction of the sequence up to \(r - 1 < L\) and construct the next execution \(\alpha_r\lambda_r\phi_r\) as follows.

By the induction hypothesis, the execution \(\alpha_{r-1}\lambda_{r-1}\) is \((r-1)\)-perturbing. If \(\alpha_{r-1}\) is saturated, we take \(\alpha_r = \alpha_{r-1}\) and \(\lambda_r = \lambda_{r-1}\). Otherwise, by case (2) of Definition 2, there is a process \(p_{\ell_r} \neq p_n\) that does not take steps in \(\lambda_{r-1}\), for which there is an extension of \(\alpha_{r-1}, \gamma_r\), consisting of events by \(p_{\ell_r}\), such that \(p_n\) returns different responses when performing \(op_n\) by itself after \(\alpha_{r-1}\lambda_{r-1}\) and after \(\alpha_{r-1}\gamma_r\lambda_{r-1}\). As per Definition 2, let \(\gamma_r = \gamma_r' e_r \gamma_r''\), where \(e_r\) is the first event of \(\gamma\) such that \(op_n\) returns different responses after \(\alpha_{r-1}\lambda_{r-1}\) and after \(\alpha_{r-1}\gamma_r' e_r \lambda_{r-1}\). Clearly \(e_r\) is a nontrivial event.

Denote by \(\phi_r\) the execution of \(op_n\) by \(p_n\) after \(\alpha_{r-1}\gamma_r' e_r \lambda_{r-1}\). Since \(op_n\) returns different values after \(\alpha_{r-1}\lambda_{r-1}\) and after \(\alpha_{r-1}\gamma_r' e_r \lambda_{r-1}\), and since the implementation uses only historyless primitives, this implies that \(e_r\) is applied to some base object \(B\) not in \(\{O_1, \ldots, O_{r-1}\}\) that is accesses by \(p_n\) in \(\phi_r\).

Define \(\alpha_r = \alpha_{r-1}\gamma_r'\) and \(\lambda_r = \lambda_{r-1} e_r\). To conclude the proof, we need to show that the execution \(\alpha_r\lambda_r\) satisfies the properties of Definition 5. Since Property 1 holds for execution \(\alpha_{r-1}\lambda_{r-1}\), it is \(p_{\ell_r}\)-free. By construction, \(\gamma_r'\) is performed by \(p_{\ell_r} \neq p_n\), hence \(\alpha_r\lambda_r\) is also \(p_{\ell_r}\)-free, establishing that Property 1 holds for it as well. Property 4 holds for \(\alpha_r\lambda_r\) since \(|active(\alpha_r\lambda_r)| = |active(\alpha_{r-1}\lambda_{r-1})| + 1 \leq r - 1 + 1 = r\). Finally, Properties 3 and 2 are immediate from our construction.

By its construction, \(\alpha_r\lambda_r\) is \(r\)-perturbing, which concludes the proof. \(\square\)
Theorem 11 The space complexity of any obstruction-free implementation of an m-limited-use max register or an m-limited-use collect object from historyless primitives is \(\Omega(\min(m, n))\). The space complexity of any obstruction-free implementation of a b-valued compare-and-swap object from historyless primitives is \(\Omega(\min(\sqrt{m}, n))\). The space complexity of any obstruction-free implementation of a k-additive-accurate m-limited-use counter from historyless primitives is \(\Omega(\min(\sqrt{m/n}, n))\).

Lemma 12 An L-perturbable object implementation has an access-stall perturbation sequence of length L.

Proof: Let \(\mathcal{O}\) be an L-perturbable object implementation. We show that it has an access-stall-perturbation sequence of length L, for the operation \(op_n\) as specified in Definition 4. The proof is by induction, where we prove the existence of the execution \(\alpha_r\sigma_{r,1} \cdots \sigma_{r,j}, \rho_r\), for every \(r, 0 \leq r \leq L\).

For the base case, \(r = 0\), \(\alpha_0\) is empty, and \(j_0 = 0\), implying that \(\lambda_0\) is also empty. It follows that \(\alpha_0\lambda_0\) is the empty execution and therefore, by Definition 2, is 0-perturbing. We next assume the construction of the sequence up to \(r < L\) and construct the next access-stall execution, \(\alpha_{r+1}\sigma_{r+1,1} \cdots \sigma_{r+1,j+1}, \rho_{r+1}\).

By induction hypothesis, \(\alpha_r\sigma_{r,1} \cdots \sigma_{r,j}, \rho_r\) is an \(r\)-perturbing execution. If it is saturated, then we set \(\alpha_{r+1} = \alpha_r, j_{r+1} = j_r, \sigma_{r+1,j} = \sigma_{r,j}\) for \(j = 1, \ldots, j_r\) and \(\rho_{r+1} = \rho_r\). By induction hypothesis and Definitions 2 and 4, \(\alpha_{r+1}\sigma_{r+1,1} \cdots \sigma_{r+1,j+1}, \rho_{r+1}\) is an \(r+1\) access-stall execution.

Assume, then, that \(\alpha_r\sigma_{r,1} \cdots \sigma_{r,j}, \rho_r\) is not saturated. Let \(\phi_r\) denote a solo execution of \(op_n\) by \(p_n\) after \(\alpha_r\lambda_r\). Since all the events in \(\lambda_r\) are by distinct processes other than \(p_n\), and since each of the objects \(O_j^k\) is accessed by \(p_n\) after it is accessed by the processes of \(S_j^k\), for \(j \in \{1, \ldots, j_r\}\), executions \(\alpha_r\sigma_{r,1} \cdots \sigma_{r,j}, \rho_r\) and \(\alpha_r\lambda_r\phi_r\) are indistinguishable to all processes. Since \(\alpha_r\lambda_r\) is an \(r\)-perturbing execution and \(r < L\), and since \(\alpha_r\lambda_r\) is not saturated, it follows from Definition 2 that there exists a process \(p_{\ell_{r+1}} \neq p_n\) that applies no event in \(\lambda_r\) and an extension \(\gamma_{r+1}^{\ell_{r+1}}\) of \(\alpha_r\), consisting of events by \(p_{\ell_{r+1}}\), such that \(e_{r+1}\) is the first event of \(\gamma_{r+1}^{\ell_{r+1}}e_{r+1}\) such that \(p_n\) returns different responses after \(\alpha_r\lambda_r\) and after \(\alpha_r\gamma_{r+1}^{\ell_{r+1}}e_{r+1}\). It follows that \(e_{r+1}\) is a nontrivial event applied by \(p_{\ell_{r+1}}\) to a base object in \(\{B_1^k, \ldots, B_{r+1}^k\}\); let this base object be \(B_{r+1}^k\). There are two cases:

Case 1: If \(B_{r+1}^k = O_{r+1}^k\), for some \(k' \in \{1, \ldots, j_r\}\), then let \(j_{r+1} = k'+1\), for \(j = 1, \ldots, k'-1, \sigma_{r+1,j} = \sigma_{r,j}\) (thus, \(O_{r+1}^j = O_j^k\) and \(S_{r+1}^j = S_j^k\)), \(O_{r+1}^{j+1} = O_{r+1}^k\), \(S_{r+1}^{j+1} = S_{r+1}^k\) \(\cup \{p_{\ell_{r+1}}\}\) (thus, \(e_{r+1}\) appears in \(\sigma_{r+1,k'}\)), \(\alpha_{r+1} = \alpha_r\gamma_{r+1}^{\ell_{r+1}}\lambda_r\), and \(\lambda_{r+1} = \lambda_r''\), where \(\lambda_r''\) consists of the events of \(\lambda_r\) applied to objects \(O_r^{k+1}, \ldots, O_r^{j'},\) and \(\lambda_r''\) consists of the events of \(\lambda_r\) applied to objects \(O_1^k, \ldots, O_r^{k'}\) and the event \(e_{r+1}\).

Case 2: Otherwise, let \(k'\) be the largest integer such that \(O_{r+1}^j\) precedes \(B_{r+1}^k\) in \(\pi_r\) (or 0 if \(B_{r+1}^k\) is not preceded in \(\pi_r\) by any of the objects \(O_1^k, \ldots, O_r^{j'}\)). Then \(j_{r+1} = k'+1\), for \(j = 1, \ldots, k'-1, \sigma_{r+1,j} = \sigma_{r,j}\) (hence also \(O_{r+1}^j = O_j^k\) and \(S_{r+1}^j = S_j^k\)), \(O_{r+1}^{j+1} = B_{r+1}^k\), \(S_{r+1}^{j+1} = \{p_{\ell_{r+1}}\}\), \(\alpha_{r+1} = \alpha_r\gamma_{r+1}^{\ell_{r+1}}\lambda_r''\), and \(\lambda_{r+1} = \lambda_r''\), where \(\lambda_r''\) consists of the events of \(\lambda_r\) applied to objects \(O_r^{k+1}, \ldots, O_r^{j'}\), and \(\lambda_r''\) consists of the events of \(\lambda_r\) applied to objects \(O_1^k, \ldots, O_r^{k'}\) and the event \(e_{r+1}\). Then \(e_{r+1}\) applies \(e_{r+1}\) and finally \(p_n\) applies its first event to \(O_{r+1}^{j+1}\), then \(p_{\ell_{r+1}}\) applies \(e_{r+1}\) and finally \(p_n\) applies its first event to \(O_{r+1}^{j+1}\).

In both cases, it follows from the construction and from Definition 2 that \(\alpha_{r+1}\lambda_{r+1}\) is (r+1)-perturbing. Since \(\alpha_r\) is \(p_n\)-free and none of the events of \(\gamma_{r+1}^{\ell_{r+1}}\lambda_r''\) are by \(p_n\), \(\alpha_{r+1}\) is also \(p_n\)-free. Let \(\rho_{r+1}\) denote the execution in which process \(p_n\) runs solo after \(\alpha_{r+1}\sigma_{r,1} \cdots \sigma_{r,j+1}\) until it completes...
the operation instance \( op \), in the course of which it accesses the base objects \( B^{1}_{r+1}, \ldots, B^{i_{r+1}} \). It follows from our construction that \( \alpha_{r+1} \sigma_{r+1,1} \cdots \sigma_{r+1,j_{r+1}} \rho_{r+1} \) is an \( r + 1 \) access-stall execution. 

**Lemma 13** \( \Phi_{r+1} > \Phi_{r} \).

**Proof:**

\[
\Phi_{r+1} = \sum_{i=1}^{i_{r+1}} n_{r+1}^{i} \cdot s^{s-i} \\
= \sum_{i=1}^{k} n_{r+1}^{i} \cdot s^{s-i} \\
= \sum_{i=1}^{k-1} n_{r+1}^{i} \cdot s^{s-i} + (n_{r+1}^{k} + 1) \cdot s^{s-k} \\
> \sum_{i=1}^{k} n_{r}^{i} \cdot s^{s-i} + \sum_{i=k+1}^{s} (s-i) \cdot s^{s-i} \\
\geq \Phi_{r} 
\]

\[\square\]

**Theorem 14** An \( n \)-process obstruction-free implementation of an \( m \)-limited-use max register, \( m \)-limited-use counter, an \( m \)-limited-use \( b \)-valued compare-and-swap object or an \( m \)-limited-use collect object from any read-modify-write primitives has an operation instance that either requires \( \Omega(\min(\log m, n)) \) steps or incurs \( \Omega(\min(\log m, n)) \) stalls. An obstruction-free implementation of a \( b \)-bounded max register from any read-modify-write primitives has an operation instance that either requires \( \Omega(\min(\log b, n)) \) steps or incurs \( \Omega(\min(\log b, n)) \) stalls. An obstruction-free implementation of a \( k \)-additive-accurate \( m \)-limited-use counter from any read-modify-write primitives has an operation instance that either requires \( \Omega(\min(\log m - \log k, n)) \) steps or incurs \( \Omega(\min(\log m - \log k, n)) \) stalls.

**C Proofs Omitted from Section 4**

**Lemma 4 (repeated)** Among the schedules \( \Xi_{k} \) defined in Section 4, \( \alpha \) returns at most \( (1 + 1/p)^r \) distinct values on average, where the average is taken over the random choices of the adversary for when to delay operations.

**Proof:** For simplicity, let us assume that \( \alpha \) performs only read operations; as observed in [3], for the particular class of executions we are considering we can always replace an \( \alpha \) that does not perform only read operations by an optimized version that reads each register on its first access and replaces any subsequent accesses with internal simulations. Notice that this means we allow the algorithm to be aware of our set of restricted executions.

For each \( \Xi_{k} \), define a bit-vector \( b^{k} = b_{1}^{k}, b_{2}^{k}, \ldots b_{r}^{k} \), where each \( b_{i}^{k} \) is 1 if either there are fewer than \( i \) reads in \( \alpha \) or the \( i \)-th read observes a value that was written by some \( \delta_{j} \), and 0 if the \( i \)-th read in \( \alpha \) reads a register that was not written by some delayed operation \( \delta_{j} \).

We will compute an upper bound \( T(b^{k}) \) on the expected number of distinct sequences of values that \( \alpha \) may observe in \( \Xi_{k}, \Xi_{k+1}, \ldots \) as a function of the bit-vector \( b^{k} \). The expectation is taken over the sequence of adversary choices used to construct the parent schedule \( \Gamma \), conditioned on the choices, made in the first \( k \) steps of the construction, that contribute to \( \Xi_{k} \). Note that the count includes the sequence of values observed by \( \alpha \) in \( \Xi_{k} \), so, for example, \( T(111111) \) will be 1 (since no further changes in the bit-vector are possible) and not 0.
Moving from $\Xi_k$ to $\Xi_{k+1}$ involves adding one step, which may change at most one register. To be visible to $\alpha$, this register must not be covered in $\Xi_k$ and must appear somewhere in the sequence of registers that $\alpha$ reads. The effect of changing the $i$-th register that $\alpha$ reads is that (a) $\alpha$ sees a new sequence of values, and (b) $\alpha$ can change what registers it reads on steps $i+1, i+2, \ldots, r$. It may also be the case (with probability $p$, independent of previous choices) that $b_{k+1}^i$ becomes 1, because the new operation on the $i$-th register becomes one of the covering operations $\delta_j$.

From the point of view of the $b_k$ vector, the effect of adding an operation that changes the sequence of values seen by $\alpha$ is to (a) set $b_{k+1}^i$ to 1 with probability $p$, for some $i$ with $b_k^i = 0$; and (b) allow any later bits in $b_{k+1}^i$ to be set arbitrarily. We can bound the number of distinct sequences seen by $\alpha$ by bounding the number of times this can happen before all bits in $b_k$ are 1.

Define the following recurrence on suffixes of $b$. Here $T(x)$ represents an upper bound on the expected number of distinct sequences of values that can be obtained from the last $|x|$ registers read by $\alpha$, if we do not change the value of any earlier registers.

\[
T(\langle \rangle) = 1.
\]
\[
T(1x) = T(x).
\]
\[
T(0x) = T(x) + (1 - p) \max_{x'} T(0x') + p \max_{x'} T(1x').
\]

The justification for this recurrence is that (a) there is only one possible sequence of values that can be obtained by reading no registers; (b) if the first register in the suffix is covered, then any further values must be obtained by updating later registers; and (c) if the first register is not covered, then we obtain at most $T(x')$ distinct values in the tail before having to apply an operation that changes the value in the first register, after which we get $T(0x')$ more values if we are lucky enough not to cover it and $T(1x')$ more values otherwise, where $x'$ is the new tail that holds after the first register is changed.

We can eliminate the max functions in the last line by observing that setting $x'$ to the all-zero vector maximizes $T(1x')$ and $T(0x')$. This is because any sequences of values that can be obtained starting with some registers covered can be obtained using the same strategy when those registers are not covered but the adversary chooses not to use them. So we can rewrite the last line as

\[
T(0x) = T(x) + (1 - p)T(0|x|) + pT(10|x|),
\]
where $|x|$ is the length of $x$, so that $0|x|$ is the all-zero vector of the same length.

For the vector $0^\ell$, this gives the recurrence

\[
T(0^\ell) = T(0^{\ell-1}) + (1 - p)T(0^\ell) + pT(10^{\ell-1}) = T(0^{\ell-1}) + (1 - p)T(0^\ell) + pT(0^{\ell-1}),
\]
which we can solve to get

\[
T(0^\ell) = \frac{1 + pT(0^{\ell-1})}{p}.
\]

It follows that $T(0^\ell) = (1 + 1/p)^\ell$. Setting $\ell = r$ completes the proof. \hfill $\square$

**Theorem 3 (repeated)** For any randomized implementation of an $m$-valued $c$-multiplicative-accurate counter for $n \geq m$ processes, and any fixed $\epsilon > 0$, there is an oblivious adversary strategy that yields, with probability at least $1 - \epsilon$, an execution consisting of at most $m - 1$ concurrent `CounterIncrement` operation instances, some of which may be incomplete, followed by a
CounterRead operation instance, in which one of the following conditions holds: (a) a constant fraction of the CounterIncrement instances take more than \( w \) operations; (b) the value returned by the CounterRead operation instance is not consistent with any linearization of the completed operation instances; or (c) the CounterRead operation instance takes \( \Omega \left( \frac{\log \log m - \log \log c}{\log w} \right) \) operations.

Proof: Let \( p = 1/4w \), let each \( \beta_i \) be a CounterIncrement and let \( \alpha \) be a CounterRead for a \( c \)-multiplicative-accurate counter. We pick out a subfamily of executions \( \Xi_{k_0}, \Xi_{k_1}, \ldots, \Xi_{k_{t-1}} \) where we choose \( k_i \) so that the schedule \( \Xi_{k_i} \) includes the first \( m_i = \lceil (2c)^2 \sqrt{m} \rceil \) calls to CounterIncrement, where \( i \) ranges from 0 to \( \ell - 1 = \left\lfloor \frac{1}{2} \log_2 c \sqrt{m} \right\rfloor - 1 = \Theta(\log m/\log c) \). Call an execution \( \Xi_{k_i} \) good if (a) it includes at most \( m_i/4 \) CounterIncrement operations that fail to terminate after executing \( w \) operations and (b) the CounterRead operation returns a value after at most \( r \) operations that is consistent with some linearization of the execution.

Our choice of \( p = 1/4w \) implies that any CounterIncrement is truncated with probability at most 1/4. Standard Chernoff bounds then give that the number of truncated increments among the first \( m_i \) increments is at most \( m_i/4 + O(\sqrt{m_i \log m_i}) = m_i(1/4 + o(1)) \) with probability at least \( 1 - m^{-a} \) for any fixed constant \( a \) (this uses the fact that \( m_i = \Omega(\sqrt{m}) \)); this implies that the same bound holds for all \( \ell + 1 \ll n \) executions with probability at least \( 1 - n^{-\alpha+1} \). Adding these truncated increments to the at most \( m_i/4 \) increments that fail to finish gives at most \( m_i(1/2 + o(1)) < 3/4 m_i \) partial increments in each good \( \Xi_{k_i} \) with high probability for sufficiently large \( m \).

It follows that if \( v_i \) is the number of increments that complete before the CounterRead in a good execution \( \Xi_{k_i} \), we have \( m_i/4 < v_i \leq m_i \) and for the return value \( x_i \) of a correct \( c \)-multiplicative-accurate CounterRead we have \( m_i/4c < x_i \leq cm_i \). From our definition of \( m_i \), we have \( m_i = m_{i+1}/4c^2 \), so \( x_i \leq cm_i = cm_{i+1}/4c^2 = m_i/4c < x_{i+1} \); it follows that the return value of the CounterRead in any two good executions is distinct.

From Lemma 4, across all executions \( \Xi_k \) (including our chosen executions \( \Xi_{k_i} \), there are an expected \( (4w + 1)^\ell \) possible distinct return values for CounterRead operations that finish in less than \( r \) steps. Markov’s inequality gives a probability of at least 1/2 that the actual number of return values is bounded by \( 2(4w + 1)^r \). So with probability at least 1/2 – \( o(1) \), there are at most \( 2(4w + 1)^r \) good executions.

We now have the adversary choose one of the \( \ell \) executions \( \Xi_{k_i} \) uniformly at random. Taking into account the probability of errors so far, the probability that the adversary chooses a good execution is at most \( (1/2 – o(1))(4w + 1)^r/\ell = (1 – o(1))(4w + 1)^r/\ell \). Let \( (1–o(1))(4w + 1)^r/\ell = \epsilon \), and solve for \( r \) by taking the log of both sides:

\[
o(1) + r \log(4w + 1) - \log \ell \geq \log \epsilon
\]

which gives

\[
r \geq \frac{\log \ell - \log(1/\epsilon) - o(1)}{\log(4w + 1)} = \Omega \left( \frac{\log \log m - \log \log c}{\log w} \right),
\]

where the \( \log(1/\epsilon) \) and \( o(1) \) terms disappear into the constant. If \( r \) is smaller than this bound, then the proportion of good executions will drop to less than \( \epsilon \).

So far we have been considering only deterministic counter implementations. However, any randomized counter implementation can be treated as a random superposition of deterministic counter implementations, each characterized by some fixed set of coins, and our adversary strategy does not depend on the particular deterministic implementation chosen. By Yao’s Principle, the same probability of getting a bad execution applies even for randomized algorithms. \( \square \)