

Average case quantum lower bounds for computing the boolean mean

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June 2003

Abstract

We study the average case approximation of the Boolean mean by quantum algorithms. We prove general query lower bounds for classes of probability measures on the set of inputs. We pay special attention to two probabilities, where we show specific query and error lower bounds and the algorithms that achieve them. We also study the worst expected error and the average expected error of quantum algorithms and show the respective query lower bounds. Our results extend the optimality of the algorithm of Brassard et al.

1 Introduction

Quantum computers can solve certain problems significantly faster than classical computers. One of these problems is the approximation of the mean of a Boolean function or, equivalently, the approximation of the mean of n Boolean variables. Suppose that the input is presented as a black-box or an oracle, which the algorithm queries [2]. Classical algorithms require $\Theta(n(1 - \varepsilon))$ evaluations (or queries) in the worst case, for error at most ε . Classical randomized algorithms solve this problem faster by requiring $\Theta(\min\{\varepsilon^{-2}, n\})$ evaluations. Quantum algorithms solve the problem in the worst case with high probability and are superior because they require only $\Theta(\min\{\varepsilon^{-1}, n\})$ queries.

More specifically, Brassard et al.[4] exhibited an algorithm achieving accuracy ε with a number of queries proportional to $\min\{\varepsilon^{-1}, n\}$. This algorithm is based on Grover's quantum search algorithm; see [9] for a description of Grover's algorithm and for details about quantum computing. The lower bounds of Nayak and Wu [10] establish the asymptotic optimality of the algorithm of Brassard et al. in the worst case.

Instead of the worst case error, we can consider the average error of quantum algorithms with respect to a probability measure on the class of the inputs. The average case is important

for two reasons. The first one is that it may reduce the query complexity. The second is that if we know that the query complexity is not reduced then the worst case results and the optimality of known algorithms is extended. It is also important to derive classes of measures for which similar complexity results hold. In this paper we deal with these issues.

In particular, for the approximation of the mean of n Boolean variables with uniform distribution on the set of inputs, the average error of any quantum algorithm, with T queries of order $o(n)$, is $\Omega(\min\{n^{-1/2}, T^{-1}\})$. The query complexity is zero as long as ε is $\omega(n^{-1/2})$ ¹. When $\varepsilon = \Theta(n^{-1/2})$ the query complexity remains zero as long as the asymptotic constant is large, but when this constant is small the query complexity is $\Omega(n^{1/2})$. The query complexity becomes asymptotically equal to that of the worst case when ε is $o(n^{-1/2})$. On the other hand, if all possible values of the mean are uniformly distributed then the average error of any algorithm is $\Omega(T^{-1})$. In this case, the query complexity is asymptotically equal to that of the worst case for all values of ε .

We generalize our results by showing conditions on classes of measures under which the query complexity is asymptotically equal to that of the worst case as long as ε is appropriately small. Our results extend the optimality of the quantum algorithm of Brassard et al. when high accuracy is important.

Quantum algorithms are probabilistic in nature. For a given input, they can produce various outcomes, each with a certain probability. Typically, we want them to achieve a given accuracy with probability greater than $\frac{1}{2}$ and, therefore, we study their probabilistic error. On the other hand, we can study the expected error of a quantum algorithm by considering its average error with respect to all outcomes resulting from a given input. Therefore, for a class of inputs we study the worst expected error. This is also an intuitive error criterion and is similar to the way we measure the error in Monte Carlo integration. We show that any algorithm with worst expected error at most ε must make $\Omega(\varepsilon^{-1}, n)$ queries. Therefore, the algorithm of Brassard et al. with repetitions as described in [6] is asymptotically optimal.

We also show that the query lower bounds that hold for the average case remain valid when we consider the average expected error of quantum algorithms. In this case we consider a probability measure on the set of inputs, and for each input we consider the expected error of the algorithm with respect to all possible outcomes.

Finally, it is easy to see that an algorithm approximating the Boolean mean can be used to approximately count the number of ones among n Boolean variables. Therefore, all our results directly extend to approximate counting and we exhibit the corresponding query and error lower bounds.

2 Problem Definition

Let $B_n = \{0, 1\}^n$ denote all tuples of n Boolean variables. We assume that any $X = (x_1, \dots, x_n) \in B_n$ is given by an oracle or a black box, which on input i outputs x_i . Oracle

¹ $f(n)$ is $\omega(g(n)) \Leftrightarrow g(n)$ is $o(f(n))$.

access of this type is called a *query*. We want to compute the mean of X , i.e.,

$$a_X = \frac{|X|}{n}, \quad \text{with } |X| = \sum_{i=1}^n x_i.$$

In this paper we consider the quantum query model of Beals et al.[2], where the cost of an algorithm is the number of its queries. A quantum algorithm applies a sequence of unitary transformations, which include queries, to an initial state, and at the end the final state is measured. See [2, 5, 9] for the details of the model of computation, which we summarize below to the extent necessary for this paper.

A quantum algorithm has the form

$$U_T Q_X U_{T-1} Q_X \cdots U_1 Q_X U_0 |\psi_0\rangle =: |\psi\rangle,$$

where U_0, \dots, U_T are unitary transformations that do not depend on the input X , the operator Q_X is also a unitary transformation and corresponds to a query to the oracle, the integer T is the number of times Q_X is applied, that is the number of queries, $|\psi_0\rangle$ is the initial state on which the sequence of transformations is applied, and $|\psi\rangle$ is the final state of the algorithm which is measured. The states $|\psi_0\rangle$ and $|\psi\rangle$ are unit vectors of $\mathcal{H}_m = \underbrace{\mathbb{C}^2 \otimes \cdots \otimes \mathbb{C}^2}_m$,

for some appropriately chosen $m \in \mathbb{N}$. The measurement produces one of M outcomes. Outcome $j \in \{0, \dots, M-1\}$ occurs with probability $p_X(j)$, which depends on j and the input X .

In principle, quantum algorithms may have many measurements applied between sequences of unitary transformations of the form above. However, any algorithm with many measurements and a total of T queries can be simulated by an algorithm with only one measurement that has $2T$ queries [5]. Hence, without loss of generality we consider the cost of algorithms with a single measurement.

Given an outcome j , we approximate a_X by a number $\hat{a}_X(j)$. Note that $\hat{a}_X(j)$ depends on the input X and the outcome of the measurement. Given a probability $p > \frac{1}{2}$, the error of a quantum algorithm with T queries on input X is defined by

$$e(X, T, p) = \inf \left\{ \gamma : \sum_{j: |a_X - \hat{a}_X(j)| \leq \gamma} p_X(j) \geq p \right\}.$$

The *worst probabilistic* error of a quantum algorithm with T queries in the class B_n is defined by

$$e^{\text{wp}}(B_n, T, p) = \max_{X \in B_n} e(X, T, p).$$

As we have mentioned, Brassard et al.[4] show a quantum summation algorithm (QS) for computing the Boolean mean and study its properties using this error criterion. The query lower bound $\Omega(\min(\varepsilon^{-1}, n))$ of Nayak and Wu [10] also holds in the worst case.

In this paper we consider the *average probabilistic* error of a quantum algorithm in the class B_n which we define by

$$e^{\text{ap}}(B_n, T, p) = \sum_{X \in B_n} e(X, T, p) \mu(X),$$

where μ is a probability measure on the set of inputs B_n .

In the next section we exhibit conditions for classes of probability measures and prove the corresponding query lower bounds. We will pay special attention to the following two measures

$$\begin{aligned}\mu_1(X) &= 2^{-n}, \quad \forall X \in B_n \\ \mu_2(X) &= \frac{1}{(n+1)\binom{n}{k}}, \quad \text{for } X \in B_n \text{ with } |X| = k.\end{aligned}$$

The first measure corresponds to the case where all inputs are equally likely, while the second measure corresponds to the case where all possible values of the mean are equally likely.

We now define the *worst expected* error of a quantum algorithm with T queries in the class B_n as

$$\begin{aligned}e^{(q)}(X, T) &= \left\{ \sum_{j=0}^{M-1} |a_X - \hat{a}_X(j)|^q p_X(j) \right\}^{1/q} \\ e^{\text{we}}(q, B_n, T) &= \max_{X \in B_n} e^{(q)}(X, T), \quad \text{with } 1 \leq q < \infty,\end{aligned}$$

where the summation is over all possible outcomes. Note that in this case we consider all the outcomes and not just outcomes that occur with probability $p > \frac{1}{2}$.

We also consider the average with respect to the inputs of the expected error, with respect to the outcomes, of a quantum algorithm with T queries in the class B_n . We call this the *average expected* error and we define it by

$$e^{\text{ae}}(q, B_n, T) = \sum_{X \in B_n} e^{(q)}(X, T) \mu(X), \quad \text{with } 1 \leq q < \infty,$$

where μ is a probability measure on the set of inputs B_n .

Finally, Nayak and Wu [10] show query lower bounds for a number of different problems. One of them is the computation of a Δ -*approximate count*, i.e., a number \hat{t}_X such that $|t_x - \hat{t}_X| < \Delta$, where $t_X = |X| = n \cdot a_X$ for $X \in B_n$. This problem is directly related to the approximation of the Boolean mean. We study it on the average by appropriately defining the error of a quantum algorithm. In particular, we set

$$\begin{aligned}e_1(X, T, p) &= n \cdot e(X, T, p) \\ e_1^{\text{ap}}(B_n, T, p) &= n \cdot e^{\text{ap}}(B_n, T, p) \\ e_1^{\text{we}}(q, B_n, T) &= n \cdot e^{\text{we}}(q, B_n, T), \quad \text{with } 1 \leq q < \infty, \\ e_1^{\text{ae}}(q, B_n, T) &= n \cdot e^{\text{ae}}(q, B_n, T), \quad \text{with } 1 \leq q < \infty.\end{aligned}$$

Our results concerning the mean directly extend to approximate counting by setting $\Delta = n\varepsilon$.

3 Average probabilistic error

Kwas and Woźniakowski [7] show that the QS algorithm has zero worst probabilistic error when the number of its queries is greater than $\frac{3}{2}\pi n$, for $p \leq \frac{8}{\pi^2}$. Trivially, QS also has zero

average probabilistic error in that case. Therefore, we study the error of quantum algorithms when the number of their queries is of order $o(n)$.

It is convenient to deal with arbitrary measures first and then use the results in the study of μ_1 and μ_2 . So we begin by defining classes of probability measures and deriving the corresponding query lower bounds. Roughly speaking, all the measures μ in a class satisfy the same lower bound for $\mu(X)$ as long $|X|$ belongs to a certain subset I of $\{0, \dots, n\}$. This lower bound depends on $|X|$, on I and, particularly, its cardinality $|I|$. For example, for the set $I = \{\frac{n}{2} - \sqrt{n}, \dots, \frac{n}{2} + \sqrt{n}\}$, which has cardinality $|I| = 2\sqrt{n} + 1$, we can use $c[\sqrt{n} \binom{n}{|X|}]^{-1}$, (where $c > 0$ is a constant and $|X| \in I$), as the lower bound defining the class of measures. Observe, that due to Lemma 6.1 in Appendix μ_1 asymptotically satisfies this lower bound on I . Similarly, for the set $I = \{\frac{n}{4}, \dots, \frac{3n}{4}\}$, which has cardinality $|I| = \frac{n}{2} + 1$, we can use $c[n \binom{n}{|X|}]^{-1}$ (where $c > 0$, is a constant and $|X| \in I$) as the lower bound in the definition of the class of measures. Note that μ_2 asymptotically satisfies this lower bound on I . Clearly these two choices of I distinguish two classes of probability measures. The cardinality of the set I as a function of n is important in our analysis. We will assume that $|I| \rightarrow \infty$ as $n \rightarrow \infty$ as in the previous two examples. We also consider the number of queries T and the desired accuracy ε as functions of n and carry out an asymptotic analysis. In this and in the following sections the implied asymptotic constants are absolute constants.

Theorem 3.1. *Consider the approximation of the Boolean mean. Let $I \subseteq \{0, \dots, n\}$ be a set of consecutive indices, such that its cardinality $|I|$ is $\omega(1)$ as a function of n . Assume that $k(n-k)$ is $\Theta(n^2)$ for every $k \in I$. Let μ be a probability measure on B_n such that*

$$\mu(X) = \Omega(|I|^{-1}) \frac{1}{\binom{n}{|X|}}, \quad \text{for every } X \in B_n \text{ with } |X| \in I.$$

Then for any $\varepsilon > 0$ of order $o(|I|n^{-1})$, the condition $e^{\text{ap}}(B_n, T, p) \leq \varepsilon$ implies that T must be $\Omega(\min(\varepsilon^{-1}, n))$.

Proof: We will prove the Theorem for $\varepsilon \geq 1/n$. The case $\varepsilon < 1/n$ will then follow immediately.

Consider a quantum algorithm with T queries that has error $e^{\text{ap}}(B_n, T, P) \leq \varepsilon$. Using the lower bound on $\mu(X)$ in the assumption of this theorem, we have

$$\varepsilon \geq \sum_{X \in B_n} e(X, T, p) \mu(X) \geq \frac{c}{|I|} \sum_{k \in I} \frac{1}{\binom{n}{k}} \sum_{X: |X|=k} e(X, T, p),$$

where $c > 0$ is a constant.

We multiply both sides of the inequality by n and define $\Delta = c^{-1}n\varepsilon$ and $\Delta(k) = n \sum_{|X|=k} e(X, T, p) / \binom{n}{k}$ and use the Markov inequality to obtain

$$\begin{aligned} \Delta &\geq \frac{1}{|I|} \sum_{k \in I} \Delta(k) \\ &= \frac{1}{|I|} \left\{ \sum_{k: \Delta(k) < 2\Delta} \Delta(k) + \sum_{k: \Delta(k) \geq 2\Delta} \Delta(k) \right\} \\ &\geq \frac{1}{|I|} 2\Delta n_+, \end{aligned}$$

where n_+ is the number of indices for which $\Delta(k) \geq 2\Delta$. Clearly $n_+ \leq \frac{1}{2}|I|$ and, therefore, $n_- := |I| - n_+ \geq \frac{1}{2}|I|$. Thus for at least half of the indices in I we have $\Delta(k) < 2\Delta$. We define J to be the set of all these indices. Note that Δ is $o(|I|)$ because ε is $o(|I|n^{-1})$. Without loss of generality we assume that Δ is an integer, since otherwise we can replace it by its ceiling, which does not change its order of magnitude.

Now consider $k \in J$ so that $\Delta(k) < 2\Delta$ and let $\delta(X, k) = ne(X, T, p)$ for $|X| = k$. Then for $m \geq 2$, which we will further specify later, we have

$$\begin{aligned} 2\Delta &> \frac{1}{\binom{n}{k}} \sum_{|X|=k} \delta(X, k) \\ &= \frac{1}{\binom{n}{k}} \left\{ \sum_{|X|=k: \delta(X, k) < 2m\Delta} \delta(X, k) + \sum_{|X|=k: \delta(X, k) \geq 2m\Delta} \delta(X, k) \right\} \\ &\geq \frac{1}{\binom{n}{k}} 2m\Delta \tilde{n}_+, \end{aligned}$$

where \tilde{n}_+ is the number of strings X with $|X| = k$, for which $\delta(X, k) \geq 2m\Delta$. Clearly $\tilde{n}_+ \leq \binom{n}{k}m^{-1}$ and, therefore, $\tilde{n}_- := \binom{n}{k} - \tilde{n}_+ \geq (1 - m^{-1})\binom{n}{k}$. Thus for at least \tilde{n}_- many strings X we have $\delta(X, T, p) := \delta(X, k) < 2m\Delta$.

Since $|I|$ is $\omega(1)$ we claim that for sufficiently large n there exist $k_1, k_2 \in J$ that are at least $4m\Delta$ apart whose distance does not exceed $O(\Delta)$, i.e., $|k_1 - k_2| \geq 4m\Delta$ and $|k_1 - k_2|$ is $O(\Delta)$. Indeed, if we assume otherwise, we have that $|k_1 - k_2| < 4m\Delta$ or $|k_1 - k_2|$ is $\omega(\Delta)$. Consider the indices in I in ascending order. Let $i_1 \in I$ be the first index that belongs to the set J (recall that at least half of the indices in I belong to J). Based on our assumption, the next index (greater than i_1) that belongs to J is either at a distance less than $4m\Delta$ away from i_1 or at a distance $\omega(\Delta)$ away from i_1 . We group together all the indices that are at a distance less than $4m\Delta$ away from i_1 . Clearly there are no more than $4m\Delta$ indices in the group. Now we consider the first index $i_2 \in I$ that belongs to J and is at a distance $\omega(\Delta)$ away from i_1 . We repeat the same procedure using i_2 in the place of i_1 , and we form a second group of indices that belong to J and are at a distance less than $4m\Delta$ away from i_2 . As we iterate this procedure we form groups of indices that belong to J , where each group is at a distance $\omega(\Delta)$ away from the group before it. We stop when we exhaust the indices in J . It is clear that between every two groups we have $\omega(\Delta)$ elements of I that do not belong to J . It is also clear that we have to repeat the above procedure at least $|J|/(4m\Delta) \geq |I|/(8m\Delta)$ times in order to exhaust the indices in J . Considering the indices between the consecutive groups that do not belong to J we conclude that the cardinality of I must be at least $\omega(\Delta)[|I|/(8m\Delta)] = \omega(|I|)$, which is a contradiction.

Now we use the algorithm that approximates the mean to derive another algorithm that approximates the partial Boolean function

$$f_{k_1, k_2}(X) = \begin{cases} 1 & \text{if } |X| = k_1 \\ 0 & \text{if } |X| = k_2, \end{cases}$$

where, without loss of generality, we can assume that $k_1 > k_2$.

The description of the new algorithm A is as follows: On input X , where $|X| = k_1$ or k_2 , we run the algorithm that approximates the mean and if the value of the result $\hat{a}_X(j)$ satisfies $|k_1 - n\hat{a}_X(j)| < 2m\Delta$ then the new algorithm outputs 1. It outputs 0 otherwise.

Let's look at the success probability $\Pr\{A(X) = f_{k_1, k_2}(X)\}$ of the new algorithm for the different inputs for which f_{k_1, k_2} is defined. If $|X| = k_1$ we have

$$\begin{aligned} \sum_{|X|=k_1} \Pr\{A(X) = f_{k_1, k_2}(X)\} &= \sum_{|X|=k_1} \Pr\{A(X) = 1\} \\ &\geq \sum_{|X|=k_1, \delta(X, T, p) < 2m\Delta} \Pr\{A(X) = 1\} \\ &= \sum_{|X|=k_1, \delta(X, T, p) < 2m\Delta} \Pr\{|k_1 - n\hat{a}_X(j)| < 2m\Delta\} \\ &\geq (1 - m^{-1}) \binom{n}{k_1} p, \end{aligned}$$

because $\delta(X, T, p) < 2m\Delta$ is equivalent to $e(X, T, p) < 2m\Delta n^{-1}$, which implies that $|k_1/n - \hat{a}_X(j)| < 2m\Delta n^{-1}$ holds with probability at least p . The fact that $\tilde{n}_- \geq (1 - m^{-1}) \binom{n}{k_1}$ yields the final inequality. Therefore, the probability that algorithm A fails on any input X for which $|X| = k_1$ satisfies

$$\sum_{|X|=k_1} \Pr\{A(X) \neq f_{k_1, k_2}(X)\} = \sum_{|X|=k_1} \Pr\{A(X) = 0\} \leq \binom{n}{k_1} \left(1 - p + \frac{p}{m}\right). \quad (1)$$

Let $c_2 = 1 - p + p/m$. We choose m in a way that that $c_2 < \frac{1}{2}$.

From [2] we know that the acceptance probability $q(X) = \Pr\{A(X) = 1\}$ of a quantum algorithm A is a real multilinear polynomial of degree at most $2T$, where T is the number of its queries. Recall that the symmetrization of q is the polynomial

$$q^{\text{sym}}(X) = \frac{\sum_{\pi} q(x_{\pi(1)}, \dots, x_{\pi(n)})}{n!}, \quad X = (x_1, \dots, x_n) \in B_n, \quad (2)$$

where the sum is over all permutations of the integers $1, \dots, n$. Minsky and Papert [8] show that there is a representation of q^{sym} as a univariate polynomial in $|X|$ of degree at most that of q^{sym} . For simplicity, with a slight abuse of notation we denote this univariate polynomial using the same symbol, i.e., $q^{\text{sym}}(|X|)$.

In particular, for $|X| = k_1$ we have $q(X) = 1 - \Pr\{A(X) = 0\}$, which implies $\Pr\{A(X) = 0\} = 1 - q(X) = f_{k_1, k_2}(X) - q(X)$. Thus

$$\begin{aligned} \sum_{|X|=k_1} \Pr\{A(X) = 0\} &= \binom{n}{k_1} - \sum_{|X|=k_1} q(X) \\ &= \binom{n}{k_1} - \binom{n}{k_1} q^{\text{sym}}(|Y|) \\ &= \binom{n}{k_1} [f_{k_1, k_2}(Y) - q^{\text{sym}}(|Y|)], \quad \forall |Y| = k_1, Y \in B_n. \end{aligned}$$

The second equality holds because when X has k_1 ones in particular locations and $n - k_1$ zeros in the remaining locations then the $k_1!(n - k_1)!$ permutations of X (when only the k_1 ones or the $n - k_1$ zeros are permuted) yield tuples that are identical to X . Therefore, every term in $\sum_{|X|=k_1} q(X)$ appears $k_1!(n - k_1)!$ times in the $\sum_{\pi} q(x_{\pi(1)}, \dots, x_{\pi(n)})$ of all permutations. Thus considering (2) we have $\sum_{\pi} q(x_{\pi(1)}, \dots, x_{\pi(n)}) = k_1!(n - k_1)! \sum_{|X|=k_1} q(X)$.

Using (1) and the last equality concerning the probability of failure of A we obtain

$$c_2 \geq f_{k_1, k_2}(X) - q^{\text{sym}}(|X|), \quad \forall |X| = k_1. \quad (3)$$

We work similarly when $|X| = k_2$. We have

$$\begin{aligned} \sum_{|X|=k_2} \Pr\{A(X) = f_{k_1, k_2}(X)\} &= \sum_{|X|=k_2} \Pr\{A(X) = 0\} \\ &\geq \sum_{|X|=k_2, \delta(X, T, p) < 2m\Delta} \Pr\{A(X) = 0\} \\ &= \sum_{|X|=k_2, \delta(X, T, p) < 2m\Delta} \Pr\{|k_1 - n\hat{a}_X(j)| \geq 2m\Delta\} \\ &\geq \sum_{|X|=k_2, \delta(X, T, p) < 2m\Delta} \Pr\{|k_2 - n\hat{a}_X(j)| < 2m\Delta\} \\ &\geq \left(1 - \frac{1}{m}\right) \binom{n}{k_2} p, \end{aligned}$$

because $\delta(X, T, p) < 2m\Delta$ is equivalent to $e(X, T, p) < 2m\Delta n^{-1}$, which implies that $|k_2/n - \hat{a}_X(j)| < 2m\Delta n^{-1}$ holds with probability at least p . The fact that $\tilde{n}_- \geq (1 - m^{-1}) \binom{n}{k_2}$ yields the final inequality. Therefore, the probability that algorithm A fails on any input X for which $|X| = k_2$ satisfies

$$\sum_{|X|=k_2} \Pr\{A(X) \neq f_{k_1, k_2}(X)\} = \sum_{|X|=k_2} \Pr\{A(X) = 1\} \leq \binom{n}{k_2} \left(1 - p + \frac{p}{m}\right). \quad (4)$$

In terms of $q(X)$ and its symmetrization the last inequality becomes

$$\binom{n}{k_2} c_2 \geq \sum_{|X|=k_2} q(X) = \binom{n}{k_2} q^{\text{sym}}(|Y|), \quad \forall |Y| = k_2, Y \in B_n,$$

where the inequality is obtained from (4) with $c_2 = 1 - p + p/m$, and the equality holds for the same reasons as those concerning the permutations of only ones or zeros in X which we explained before. This implies

$$c_2 \geq q^{\text{sym}}(|Y|) = |f_{k_1, k_2}(Y) - q^{\text{sym}}(|Y|)|, \quad \forall |Y| = k_2. \quad (5)$$

We combine (3) and (5) to obtain

$$|f_{k_1, k_2}(X) - q^{\text{sym}}(|X|)| \leq c_2 < \frac{1}{2},$$

for all the X for which this partial Boolean function is defined. Recall that symmetrization does not increase the degree of a polynomial, which implies that $2T$ is greater than or equal to the degree of q^{sym} . Using the results of Nayak and Wu [10] concerning lower bounds for the degree of polynomials approximating the partial Boolean function f_{k_1, k_2} , and our assumption that $k(n-k) = \Theta(n^2)$, for all $k \in I$, we obtain that the degree of q^{sym} is

$$\Omega\left(\sqrt{\frac{n}{|k_1 - k_2|}} + \frac{\sqrt{\kappa(n - \kappa)}}{|k_1 - k_2|}\right),$$

where $\kappa \in \{k_1, k_2\}$ which maximizes $|\frac{n}{2} - \kappa|$. Therefore, the number of queries of the original algorithm is $\Omega(n\Delta^{-1})$, which, in turn, is $\Omega(\varepsilon^{-1})$. \blacksquare

Theorem 3.1 extends the optimality properties of QS to the average probabilistic case when high accuracy is important. It shows that QS is asymptotically optimal in computing the Boolean mean as long as μ satisfies certain properties. The range of possible values of ε has to be appropriately small and this depends on the class of measures through the cardinality of the set I . The larger this set is the larger the range of ε for which Theorem 3.1 holds and QS is asymptotically optimal. On the other hand, as we are about to see, when there is demand for relatively low accuracy there can be other algorithms faster than QS.

Let us now consider μ_1 where all elements $X \in B_n$ are equally likely having probability 2^{-n} . Kwas and Woźniakowski [7] show that, with probability $p = 1$, the algorithm that outputs $\frac{1}{2}$ on any input without any queries at all, i.e., $T = 0$, has error

$$e^{\text{ap}}(B_n, 0, 1) = (2\pi n)^{-1/2}(1 + o(1)). \quad (6)$$

However, reducing the error further requires $\Omega(n^{-1/2})$ queries, as we see below.

Lemma 3.1. *Consider the measure μ_1 . There exists a constant $c > 0$ such that the condition $e^{\text{ap}}(B_n, T, p) \leq cn^{-1/2}$, $p > \frac{1}{2}$, implies that T is $\Omega(n^{1/2})$.*

Proof: The proof is very similar to that of Theorem 3.1. We point out the differences and we refer to the proof of Theorem 3.1 for the identical parts.

Recall that in the proof of Theorem 3.1 equation (1) lead us to select $m \geq 2$ such that $1 - p + p/m < \frac{1}{2}$. Consider any such m here.

We set $c = e^{-6(m+1)^2 - 2}(2\pi)^{-1/2}$. We consider the sets $I_1 = \{\frac{n}{2} + m\sqrt{n}, \dots, \frac{n}{2} + (m+1)\sqrt{n}\}$ and $I_2 = \{\frac{n}{2} - (m+1)\sqrt{n}, \dots, \frac{n}{2} - m\sqrt{n}\}$. Note that for the indices $k \in I_1 \cup I_2$ we have $k(n-k) = \Theta(n^2)$.

Assume that

$$cn^{-1/2} \geq \sum_{X \in B_n} e(X, T, p)\mu_1(X) \geq 2^{-n} \sum_{|X| \in I_j} e(X, T, p), \quad j = 1, 2,$$

since $\mu_1(X) = 2^{-n}$, for every $X \in B_n$. From Lemma 6.1, in the Appendix, we have that $\binom{n}{|X|}2^{-n} > cn^{-1/2}$, when $|X| \in I_1 \cup I_2$, $X \in B_n$. We multiply by n both sides of the inequality above, and define $\Delta = cn^{1/2}$ and $\Delta(k) = n \sum_{|X|=k} e(X, T, p) / \binom{n}{k}$ to obtain

$$\Delta > n^{-1/2} \sum_{k \in I_j} \Delta(k), \quad j = 1, 2.$$

Thus, there exist $k_j \in I_j$, such that $\Delta(k_j) < \Delta$, $j = 1, 2$. Let $\delta(X, k_j) = ne(X, T, p)$, $|X| = k_j$, $j = 1, 2$. Then we have

$$\begin{aligned} \Delta &> \frac{1}{\binom{n}{k_j}} \sum_{|X|=k_j} \delta(X, k_j) \\ &= \frac{1}{\binom{n}{k_j}} \left\{ \sum_{|X|=k_j: \delta(X, k_j) < m\Delta} \delta(X, k_j) + \sum_{|X|=k_j: \delta(X, k_j) \geq m\Delta} \delta(X, k_j) \right\} \\ &\geq \frac{1}{\binom{n}{k_j}} m\Delta \tilde{n}_{j,+}, \quad j = 1, 2, \end{aligned}$$

where $\tilde{n}_{j,+}$ is the number of strings X with $|X| = k_j$, for which $\delta(X, k_j) \geq m\Delta$. Just like in the proof of Theorem 3.1 we conclude that the number of strings X for which $\delta(X, k_j) < m\Delta$ satisfies $\tilde{n}_{j,-} \geq (1 - m^{-1})\binom{n}{k_j}$, $j = 1, 2$.

Now we use the algorithm that approximates the mean to derive another algorithm that approximates the partial Boolean function

$$f_{k_1, k_2}(X) = \begin{cases} 1 & \text{if } |X| = k_1 \\ 0 & \text{if } |X| = k_2. \end{cases}$$

From this point on the proof is identical to the proof of Theorem 3.1. and we omit the details. The conclusion is that the algorithm that approximates f_{k_1, k_2} and, therefore, the original algorithm must make $\Omega(n/\Delta)$ or, equivalently, $\Omega(n^{1/2})$ queries. \blacksquare

Thus we need to study the error of the algorithm when the number of queries is $\Omega(n^{1/2})$. The following theorem deals with this case and also summarizes our results with respect to μ_1 .

Theorem 3.2. *Consider that approximation of the Boolean mean and the average probabilistic error of a quantum algorithm with respect to μ_1 . The following two statements hold.*

1. *Let T be $o(n)$. Then the error of any quantum algorithm with T queries satisfies*

$$e^{\text{ap}}(B_n, T, p) = \Omega(\min\{n^{-1/2}, T^{-1}\}).$$

2. *Let $\varepsilon > 0$ be $o(n^{-1/2})$. Then the number of queries $T(\varepsilon)$ for error at most ε satisfies*

$$T(\varepsilon) = \Omega(\min\{\varepsilon^{-1}, n\}).$$

Proof: The second statement directly follows from Theorem 3.1. Indeed, in the proof of Lemma 3.1 we saw a lower bound for $\mu_1(X)$ when $|X|$ belongs to sets of $n^{1/2}$ many indices close to $n/2$ (sets like I_1 and I_2). Thus the conditions of Theorem 3.1 hold for μ_1 and the query lower bound is immediate.

Now we prove the first statement. From Lemma 3.1 we know that error less than $cn^{-1/2}$ requires $\Omega(n^{1/2})$ queries. Hence, when the number of queries is $o(n^{1/2})$ then the error is bounded from below by a quantity proportional to $n^{-1/2}$.

Let us now consider T to be $\Omega(n^{1/2})$. Then T^{-1} is $O(n^{-1/2})$ and $\min\{n^{-1/2}, T^{-1}\} = \Theta(T^{-1})$. We prove the first statement by contradiction. Assume that n is sufficiently large. Suppose that the error lower bound is not $\Omega(\min\{n^{-1/2}, T^{-1}\})$ but that $e^{\text{ap}}(B_n, T, p) \leq (Tg(T))^{-1}$, where g is a function such that $g(T) = \omega(1)$. Set $\varepsilon = (Tg(T))^{-1}$ and observe that ε is $o(n^{-1/2})$. Then use the second statement of this theorem conclude that T must be $\Omega(Tg(T))$, which is a contradiction. ■

Kwas and Woźniakowski [7] show that for μ_1 , the average probabilistic error of QS is $O(\min\{n^{-1/2}, T^{-1}\})$ when the number of its queries is divisible by four. Using Theorem 3.2 we conclude:

- QS is an asymptotically optimal error algorithm.
- QS makes an asymptotically optimal number of queries for accuracy ε , when ε is $\omega(n^{-1/2})$.
- QS requires at least four queries for error $O(n^{-1/2})$ when ε is $\omega(n^{-1/2})$, while the optimal number of queries is zero, and is achieved by a constant algorithm.

We now consider μ_2 which corresponds to the case that all values of the mean are equally likely. As we shall see, computing the mean in the average probabilistic case with μ_2 is just as hard as computing the mean in the worst probabilistic case.

Theorem 3.3. *Consider the approximation of the Boolean mean and the average probabilistic error of a quantum algorithm with respect to μ_2 . The following two statements hold.*

1. *Let T be $o(n)$. Then the error of any quantum algorithm with T queries satisfies*

$$e^{\text{ap}}(B_n, T, p) = \Omega(T^{-1}).$$

2. *Let $\varepsilon > 0$ be $o(1)$. Then the number of queries $T(\varepsilon)$ for error at most ε satisfies*

$$T(\varepsilon) = \Omega(\min\{\varepsilon^{-1}, n\}).$$

Proof: Trivially μ_2 satisfies the conditions of Theorem 3.1 for a set I of $\Theta(n)$ many consecutive indices, e.g., $I = \{\frac{n}{4}, \dots, \frac{3}{4}n\}$. Therefore, the second statement is immediate.

We show the first statement by contradiction. If T is $O(1)$ then the error is bounded from below by a constant. Indeed, if we assume that $e^{\text{ap}}(B_n, T, p) \leq 1/g(n)$ for some function g satisfying $g(n) = \omega(1)$, then Theorem 3.1 yields that $T = \Omega(g(n))$, which is a contradiction. In contrast to μ_1 , the measure μ_2 does not make the problem easier.

When T is $\omega(1)$, suppose that $e^{\text{ap}}(B_n, T, p)$ is $o(T^{-1})$. Let n be sufficiently large. Then there exists a function g with $g(T) = \omega(1)$ such that $e^{\text{ap}}(B_n, T, p) \leq (Tg(T))^{-1}$. Set $\varepsilon = (Tg(T))^{-1}$ and observe that $\varepsilon = o(1)$, as the second statement of the theorem requires. This leads us to conclude that T must be $\Omega(Tg(T))$ and, therefore, we get a contradiction. ■

For μ_2 , Theorem 3.3 and the results of [4] and [7] (for the worst probabilistic error of QS) imply that QS is an asymptotically optimal error and query algorithm. Hence, in terms of

error and number of necessary queries, computing the Boolean mean on the average with μ_2 is as difficult as in the worst probabilistic case.

We end this section by extending our results to Δ -approximate count. We present three corollaries. We omit their proofs since they are immediate from the corresponding theorems above.

Corollary 3.1. *Consider Δ -approximate count. Let $I \subseteq \{0, \dots, n\}$ be a set of indices, such that its cardinality $|I|$, as a function of n , is $\omega(1)$, and $k(n-k)$ is $\Theta(n^2)$ for every $k \in I$. Assume that μ is a probability measure on B_n such that*

$$\mu(X) = \Omega(|I|^{-1}) \frac{1}{\binom{n}{|X|}}, \quad \text{for every } |X| \in I, \quad X \in B_n.$$

Then for any $\Delta > 0$ of order $o(|I|)$, $e_1^{\text{ap}}(B_n, T, p) \leq \Delta$ implies that $T = \Omega(\min(n/\Delta, n))$.

Corollary 3.2. *Consider Δ -approximate count and the average probabilistic error of a quantum algorithm with respect to μ_1 . The following two statements hold.*

1. *Let T be $o(n)$. Then the error of any quantum algorithm with T queries satisfies*

$$e_1^{\text{ap}}(B_n, T, p) = \Omega(\min\{n^{1/2}, n/T\}).$$

2. *Let $\Delta > 0$ be $o(n^{1/2})$. Then the number of queries $T(\Delta)$ for error at most Δ satisfies*

$$T(\Delta) = \Omega(\min\{n/\Delta, n\}).$$

Corollary 3.3. *Consider Δ -approximate count and the average probabilistic error of a quantum algorithm with respect to μ_2 . The following two statements hold.*

1. *Let T be $o(n)$. Then the error of any quantum algorithm with T queries satisfies*

$$e_1^{\text{ap}}(B_n, T, p) = \Omega(n/T).$$

2. *Let $\Delta > 0$ be $o(n)$. Then the number of queries $T(\Delta)$ for error at most Δ satisfies*

$$T(\Delta) = \Omega(\min\{n/\Delta, n\}).$$

4 Worst expected error

In this section we consider quantum algorithms with a worst expected error criterion. We show query lower bounds for any quantum algorithm computing the Boolean mean and for any quantum algorithm computing a Δ -approximate count.

Theorem 4.1. Consider any algorithm that computes the Boolean mean with worst expected error satisfying $e^{\text{we}}(q, B_n, T) \leq \varepsilon$, for a fixed $q \in [1, \infty)$. Then the number of queries of this algorithm satisfies

$$T(\varepsilon) = \Omega(\min\{\varepsilon^{-1}, n\}).$$

Proof: Consider $e^{\text{we}}(q, B_n, T) \leq \varepsilon$ and raise both sides to the power q and multiply them by n^q . Then set $\Delta = n\varepsilon$ and $\Delta(X, j) = n|a_X - \hat{a}_X(j)|$ to obtain

$$\Delta^q \geq \sum_{j=0}^{M-1} \Delta(X, j)^q p_X(j), \quad \forall X \in B_n.$$

For any $\delta > 0$ using the Markov inequality we have

$$\Delta^q \geq \delta^q \sum_{\Delta(X, j) \geq \delta} p_X(j), \quad \forall X \in B_n.$$

Choose a number $a > 2$ and set $\delta = a\Delta$ and $p = 1 - a^{-q}$, $p \in (\frac{1}{2}, 1)$. Define $p_{\text{loss}}(X) = \sum_{\Delta(X, j) \geq \delta} p_X(j)$ for $X \in B_n$. Then

$$1 - p = a^{-q} = \frac{\Delta^q}{\delta^q} \geq p_{\text{loss}}(X).$$

This implies that $p_{\text{win}}(X) = 1 - p_{\text{loss}}(X) \geq p > \frac{1}{2}$, $\forall X \in B_n$.

Then there exist outcomes j for which $\delta > \Delta(X, j)$ with probability

$$p_{\text{win}}(X) = \sum_{n|a_X - \hat{a}_X(j)| < \delta} p_X(j) \geq p > 1/2, \quad \forall X \in B_n.$$

Hence, the probabilistic error of δ -approximate count is $e_1(X, T, p) < \delta$, for all $X \in B_n$. Now take any X such that $|X|(n - |X|) = \Theta(n^2)$ and use the results of [10] to see that the number of necessary queries is

$$\Omega\left(\sqrt{\frac{n}{\delta}} + \frac{\sqrt{|X|(n - |X|)}}{\delta}\right).$$

Therefore, the number of queries satisfies

$$\Omega(\min\{\varepsilon^{-1}, n\}).$$

■

It has been recently shown in [6] that if one repeats $2(\lceil q \rceil + 1)$ times the QS algorithm of Brassard et al. (with T queries) then the median of the outputs has worst expected error of order $O(T^{-1})$. Using the theorem above we conclude that this is an asymptotically optimal algorithm.

The following query lower bound for Δ -approximate count is a direct consequence of Theorem 4.1.

Corollary 4.1. Consider any algorithm that computes a Δ -approximate count with worst expected error satisfying $e_1^{\text{we}}(q, B_n, T) \leq \Delta$, for fixed $q \in [1, \infty)$. Then the number of queries of this algorithm satisfies

$$T(\Delta) = \Omega(\min\{n/\Delta, n\}).$$

5 Average expected error

In this section we consider the average expected error of quantum algorithms. For brevity we call this the average expected setting. Recall that we are considering the average with respect to a probability measure on the set of inputs B_n and for each of the inputs we consider the expected error of the quantum algorithm with respect to all possible outcomes. We show query lower bounds for a quantum algorithm computing the Boolean mean and for any quantum algorithm computing a Δ -approximate count.

We deal only with the measures of Theorem 3.1 since μ_1 and μ_2 are special cases that can be dealt with in the same way. In fact, Theorem 3.1 holds for the average expected error as well. The proof is based on that of Theorem 3.1.

Theorem 5.1. *Consider the approximation of the Boolean mean. Let $I \subseteq \{0, \dots, n\}$ be a set of consecutive indices, such that its cardinality $|I|$, as a function of n , is $\omega(1)$. Assume that $k(n-k)$ is $\Theta(n^2)$ for every $k \in I$. Let μ be a probability measure on B_n such that*

$$\mu(X) = \Omega(|I|^{-1}) \frac{1}{\binom{n}{|X|}}, \quad \text{for every } X \in B_n \text{ with } |X| \in I.$$

Consider a fixed $q \in [1, \infty)$. Then for any $\varepsilon > 0$ of order $o(|I|n^{-1})$, the condition $e^{\text{ae}}(q, B_n, T) \leq \varepsilon$ implies that T must be $\Omega(\min(\varepsilon^{-1}, n))$.

Proof: The proof is almost identical to that of Theorem 3.1 and we will only point out the differences.

In particular, for $1 \leq q < \infty$ consider a quantum algorithm with average expected error at most ε , i.e.,

$$\varepsilon \geq e^{\text{ae}}(B_n, T) = \sum_{X \in B_n} e^{(q)}(X, T) \mu(X).$$

We follow the first part of the proof of Theorem 3.1 replacing $e(X, T, p)$ by $e^{(q)}(X, T)$ and redefining the rest of the quantities accordingly. After the two applications of the Markov inequality we know that the number of strings X for which $\delta(X, k) < 2m\Delta$, $|X| = k$, is $\tilde{n}_- \geq (1 - m^{-1}) \binom{n}{k}$, and $m \geq 2$. Recall that $\delta(X, k) = ne^{(q)}(X, T)$.

Using the Markov inequality as in Theorem 4.1 to derive the probabilistic error from the expected error, we conclude that the probabilistic error of approximate count satisfies $e_1(X, T, p) < 2ma\Delta$, with probability $p \geq 1 - a^{-q} > 1/2$ for a chosen $a > 2$.

Now we return to the proof of Theorem 3.1. We have that there exist $k_1, k_2 \in J$ that are at least $4ma\Delta$ apart whose distance does not exceed $O(\Delta)$, i.e., $|k_1 - k_2| \geq 4ma\Delta$ and $|k_1 - k_2| = O(\Delta)$, and $e_1(X, T, p) \leq 2ma\Delta$, $|X| = k_1$, or k_2 .

We use the original algorithm that approximates the mean to derive a new algorithm that approximates the partial Boolean function

$$f_{k_1, k_2}(X) = \begin{cases} 1 & \text{if } |X| = k_1 \\ 0 & \text{if } |X| = k_2. \end{cases}$$

where, without loss of generality, we can assume that $k_1 > k_2$.

The description of the new algorithm A is as follows: On input X , where $|X| = k_1$ or k_2 , we run the algorithm that approximates the mean and if the value of the result $\hat{a}_X(j)$ satisfies $|k_1 - n\hat{a}_X(j)| < 2ma\Delta$ then the new algorithm outputs 1. It outputs 0 otherwise.

There is one more difference between this proof and the proof of Theorem 3.1. It concerns the derivation of the success/failure probability of the new algorithm and we explain this difference below.

Let's look at the success probability $\Pr\{A(X) = f_{k_1, k_2}(X)\}$ of the new algorithm for the different inputs for which f_{k_1, k_2} is defined. If $|X| = k_1$ we have

$$\begin{aligned} \sum_{|X|=k_1} \Pr\{A(X) = f_{k_1, k_2}(X)\} &= \sum_{|X|=k_1} \Pr\{A(X) = 1\} \\ &\geq \sum_{|X|=k_1, \delta(X, k) < 2m\Delta} \Pr\{A(X) = 1\} \\ &= \sum_{|X|=k_1, \delta(X, k) < 2m\Delta} \Pr\{|k_1 - n\hat{a}_X(j)| < 2ma\Delta\} \\ &\geq (1 - m^{-1}) \binom{n}{k_1} (1 - a^{-q}), \end{aligned}$$

because we saw that when the expected error satisfies $\delta(X, k) < 2m\Delta$ then this implies that the probabilistic error satisfies $e_1(X, T, p) < 2ma\Delta$ with probability $p \geq 1 - a^{-q} > 1/2$. The fact that $\tilde{n}_- \geq (1 - m^{-1}) \binom{n}{k_1}$ yields the final inequality. Therefore, the probability that algorithm A fails on any input X for which $|X| = k_1$ satisfies

$$\sum_{|X|=k_1} \Pr\{A(X) \neq f_{k_1, k_2}(X)\} = \sum_{|X|=k_1} \Pr\{A(X) = 0\} \leq \binom{n}{k_1} \left(a^{-q} + \frac{1 - a^{-q}}{m} \right).$$

Let $c_2 = a^{-q} + (1 - a^{-q})/m$, where $a > 2$. Just like in the proof of Theorem 3.1, we choose m in a way that $c_2 < \frac{1}{2}$. This leads us to the equivalent of (1) of Theorem 3.1.

In the same way we derive the equation concerning the probability of failure of the new algorithm on input $|X| = k_2$ which corresponds to equation (4) of Theorem 3.1.

The remaining steps are identical to those of Theorem 3.1 and complete the proof. \blacksquare

Theorem 5.1 shows that QS algorithm with repetitions [6] is asymptotically optimal in the average expected case when the required accuracy is high. In fact, the query lower bounds of section 3 that depend either on Theorem 3.1 directly or have been derived through a similar proof technique extend to the average expected and we have seen how this can be accomplished in the proof of Theorem 5.1.

The following corollary for Δ -approximate count in the average expected case is immediate.

Corollary 5.1. *Consider Δ -approximate count. Let $I \subseteq \{0, \dots, n\}$ be a set of consecutive indices, such that its cardinality $|I|$, as a function of n , is $\omega(1)$, and $k(n - k)$ is $\Theta(n^2)$ for every $k \in I$. Assume that μ is a probability measure on B_n such that*

$$\mu(X) = \Omega(|I|^{-1}) \frac{1}{\binom{n}{|X|}}, \quad \text{for every } |X| \in I, \quad X \in B_n.$$

Consider a fixed $q \in [1, \infty)$. Then for any $\Delta > 0$ of order $o(|I|)$, $e_1^{\text{ae}}(q, B_n, T) \leq \Delta$ implies that $T = \Omega(\min(n/\Delta, n))$.

Acknowledgements

I thank P. Jaksch, J. Traub, A. Werschulz and H. Woźniakowski for their comments and suggestions that significantly improved this paper.

6 Appendix

Lemma 6.1. For $n \in \mathbb{N}$ and $1 \leq c \leq \sqrt{n}/6$ we have

$$\binom{n}{n/2 \pm c\sqrt{n}} > e^{-6c^2-2} \frac{2^n}{\sqrt{2\pi n}}.$$

Proof: From Stirling's formula [1, p. 257] we have

$$n! = \sqrt{2\pi n} n^{n+1/2} e^{-n+\theta/(12n)}, \quad 0 < \theta < 1.$$

Thus,

$$n! \leq e\sqrt{2\pi n}(n/e)^n,$$

and

$$n! \geq \sqrt{2\pi n}(n/e)^n.$$

Therefore,

$$\begin{aligned} (n/2 + c\sqrt{n})! &\leq e \left(\frac{n + 2c\sqrt{n}}{2e} \right)^{n/2+c\sqrt{n}} \sqrt{2\pi(n/2 + c\sqrt{n})} \\ &< e \left(\frac{n + 2c\sqrt{n}}{2e} \right)^{n/2+c\sqrt{n}} \sqrt{2\pi n}, \end{aligned}$$

and

$$\begin{aligned} (n/2 - c\sqrt{n})! &\leq e \left(\frac{n - 2c\sqrt{n}}{2e} \right)^{n/2-c\sqrt{n}} \sqrt{2\pi(n/2 - c\sqrt{n})} \\ &< e \left(\frac{n - 2c\sqrt{n}}{2e} \right)^{n/2-c\sqrt{n}} \sqrt{2\pi n}. \end{aligned}$$

From the inequalities above we obtain

$$\begin{aligned}
\frac{n!}{(n/2 + c\sqrt{n})!(n/2 - c\sqrt{n})!} &> \frac{2^n n^n}{e^2 \sqrt{2\pi n} (n + 2c\sqrt{n})^{n/2 + c\sqrt{n}} (n - 2c\sqrt{n})^{n/2 - c\sqrt{n}}} \\
&= \frac{2^n n^n}{e^2 \sqrt{2\pi n} (n + 2c\sqrt{n})^{n/2} (n - 2c\sqrt{n})^{n/2}} \left(\frac{n - 2c\sqrt{n}}{n + 2c\sqrt{n}} \right)^{c\sqrt{n}} \\
&= \frac{2^n n^n}{e^2 \sqrt{2\pi n} (n^2 - 4c^2 n)^{n/2}} \left(\frac{\sqrt{n} - 2c}{\sqrt{n} + 2c} \right)^{c\sqrt{n}} \\
&> \frac{2^n n^n}{e^2 \sqrt{2\pi n} n^n} \left(\frac{\sqrt{n} - 2c}{\sqrt{n} + 2c} \right)^{c\sqrt{n}} \\
&= \frac{2^n}{e^2 \sqrt{2\pi n}} \left(\frac{\sqrt{n} - 2c}{\sqrt{n} + 2c} \right)^{c\sqrt{n}}
\end{aligned}$$

Using

$$\left(\frac{\sqrt{n} + 2c}{\sqrt{n} - 2c} \right)^{\sqrt{n} - 2c} = \left(1 + \frac{4c}{\sqrt{n} - 2c} \right)^{\sqrt{n} - 2c} < e^{4c},$$

we obtain that

$$\left(\frac{\sqrt{n} + 2c}{\sqrt{n} - 2c} \right)^{c\sqrt{n}} < e^{4c^2} \left(\frac{\sqrt{n} + 2c}{\sqrt{n} - 2c} \right)^{2c^2} \leq e^{6c^2}.$$

Thus,

$$\frac{n!}{(n/2 + c\sqrt{n})!(n/2 - c\sqrt{n})!} > \frac{2^n}{e^{6c^2 + 2} \sqrt{2\pi n}}.$$

■

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