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Quantum Speed-ups for Single-machine Scheduling Problems

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ABSTRACT
Grover search is currently one of the main approaches to obtain quantum speed-ups for combinatorial optimization problems. The combination of Quantum Minimum Finding (obtained from Grover search) with dynamic programming has proved particularly efficient to improve the worst-case complexity of several NP-hard optimization problems. Specifically, for these problems, the classical dynamic programming complexity (ignoring the polynomial factors) in \( O^n(c^n) \) can be reduced by a bounded-error hybrid quantum-classical algorithm to \( O^n(c_{\text{quant}}^n) \) for \( c_{\text{quant}} < c \). In this paper, we extend the resulting hybrid dynamic programming algorithm to three examples of single-machine scheduling problems: minimizing the total weighted completion time with deadlines, minimizing the total weighted completion time with precedence constraints, and minimizing the total weighted tardiness. The extension relies on the inclusion of a pseudo-polynomial term in the state space of the dynamic programming as well as an additive term in the recurrence.

1 INTRODUCTION
The interest in quantum computing to solve combinatorial optimization problems has been growing for several years in the operational research community. More precisely, two branches are distinguished. The first one relates to heuristics, often hybrid quantum-classical algorithms such as variational quantum algorithms [3, 8] and in particular QAOA [6]. Essentially, these algorithms require the optimization problem to be formulated as a QUBO (Quadratic Unconstrained Binary Optimization) and can be implemented on current noisy quantum computers because the quantum part can be made rather small. The second branch relates to exact algorithms. Unlike the previous algorithms, it is impossible to implement them today but theoretical speed-ups have been proved for several types of problems and algorithms [13, 16].

The most emblematic algorithm of this branch is Grover search [9], which achieves a quadratic speed-up when searching for a specific element in an unsorted table, where the complexity is computed as the number of queries of the table and done by an oracle. The authors of [5] use Grover search as a subroutine for a hybrid quantum-classical algorithm that finds with high probability the minimum of an unsorted table, leading to the algorithm known as Quantum Minimum Finding (QMF). Later, the authors of [2] combine QMF with dynamic programming to address NP-hard optimization problems. They apply their algorithm to vertex ordering problems, the Traveling Salesman Problem (TSP), and the Minimum Set Cover problem, among others. All these problems satisfy a specific property which implies that they can be solved by classical dynamic programming in \( \mathcal{O}(c^n) \), where \( \mathcal{O} \) is the usual asymptotic notation that ignores the polynomial factors, and \( c \) is usually not smaller than 2. The hybrid algorithm from [2] reduces the complexity to \( \mathcal{O}(c_{\text{quant}}^n) \) for \( c_{\text{quant}} < c \). For instance, the TSP is solved by Held and Karp dynamic programming [10] in \( \mathcal{O}(2^n) \), and by the hybrid algorithm of [2] in \( \mathcal{O}(1.728^n) \). Subsequently to the work of [2], other NP-hard problems have been tackled with the idea of combining Grover search (or QMF) and classical dynamic programming. This has led to quantum speed-ups for the Steiner Tree problem [11], the graph coloring problem [15], and the subset sum problem [1].

The purpose of this work is to adapt the seminal idea of [2] to NP-hard scheduling problems [17] that satisfy the following property: for a given set of jobs \( J \), the optimal solution for \( J \) is the...
best concatenation of optimal solutions for \(X\) and \(J\setminus X\) among all \(X \subseteq J\) such that \(|X| = |J|/2\) (modulo an additive term that arises in the concatenation). This adaptation requires to introduce a pseudo-polynomial term in the state space of the dynamic programming as well as the aforementioned additive term. We thus obtain an extension of the Dynamic Programming Across the Subsets (DPAS) that many scheduling problems satisfy [17]. Herein, we focus on single-machine scheduling problems and show that our bounded-error hybrid quantum-classical algorithm improves the best-known classical exponential complexities, where in some cases a pseudo-polynomial factor \(\sum p_j\) appears. We illustrate it with three examples: minimizing the total weighted completion time with deadlines, minimizing the total weighted completion time with precedence constraints, and minimizing the total weighted tardiness. We summarize in Table 1 the different complexities.

<table>
<thead>
<tr>
<th>Problem</th>
<th>Quantum-DDPAS</th>
<th>Best classical algorithm</th>
</tr>
</thead>
<tbody>
<tr>
<td>1</td>
<td>(w_j x_j)</td>
<td>(O^*(2^n)) (DPAS)</td>
</tr>
<tr>
<td>1</td>
<td>(w_j y_j)</td>
<td>(O^*(2^n)) (DPAS)</td>
</tr>
<tr>
<td>1</td>
<td>(w_j c_j)</td>
<td>(O^*(2^n)) (DPAS)</td>
</tr>
</tbody>
</table>

Table 1: Comparison of complexities between our hybrid algorithm Quantum-DDPAS and the best-known classical algorithm

The rest of the paper is structured as follows. We detail in Section 2 the required property and give examples of single-machine scheduling problems that satisfy it. Then, we describe in Section 3 the hybrid algorithm that solves the problems of interest, recalling known bounds useful to derive the complexities of the algorithm as well as the hybrid algorithm. Let \(\mathcal{P}\) be the nominal problem we want to solve. We introduce next a family of problems related to \(\mathcal{P}\) that will be instrumental in deriving the dynamic programming recursion. Let \(T\) be a set of nonnegative integers containing \(0\). We define the family of problems indexed by \(J \subseteq [n]\) and \(t \in T\):

\[
P(J, t) := \min_{\pi \in \Pi(J, t)} f(\pi, J, t),
\]

where \(\Pi(J, t) \subseteq \mathcal{S}_J\) is the set of feasible permutations of \(J\) according to potential constraints and \(f(., J, t)\) is the objective function. We note \(\text{OPT}[J, t]\) the optimal value of \(P(J, t)\). With these notations, the nominal problem \(\mathcal{P}\) can be cast as follows:

\[
\mathcal{P} = P([n], 0).
\]

### 2.1 Dynamic Programming Across the Subsets

We suppose in what follows that \(\mathcal{P}\) can be solved by DPAS (Dynamic Programming Across the Subsets). It means that the family of problems must satisfy the following DPAS property.

**Property 1 (DPAS).** Let \(t_0 \in T\). Problem \(P([n], t_0)\) can be solved by DPAS if there exists a function \(h: 2^n \times [n] \times T \rightarrow \mathbb{R}\), computable in polynomial time, such that the following holds:

\[
\text{OPT}[J, t_0] = \begin{cases} 
\min_{J \subseteq [n]} [\text{OPT}[J \setminus \{j\}, t_0] + h(J, j, t_0)] & \forall J \subseteq [n] \\
\text{OPT}([0], t_0) = 0
\end{cases}
\]

Notice the presence of the additional parameter \(t_0\) in the above definitions, which is typically absent in the scheduling literature. In particular, \(t_0\) is a constant throughout the whole recursion (2) and does not impact the resulting computational complexity. The use of that extra parameter defined in Equation (1) and in Property 1 shall be necessary later when applying our hybrid algorithm.

**Lemma 2 (DPAS complexity).** DPAS solves \(\mathcal{P}\) in \(O^*(2^n)\).

**Proof.** We compute Equation (2) for all \(J\) such that \(|J| = k\), and for \(t_0 = 0\), starting from \(k = 1\) to \(k = n\). For a given \(J\), the values \(\{\text{OPT}[J \setminus \{j\}, 0] : j \in J\}\) are known, so \(\text{OPT}[J, 0]\) is computed in time \(\text{poly}(n) \cdot k\) according to Equation (2). Eventually, the total complexity of computing \(\text{OPT}([n], 0)\) is

\[
\sum_{k=1}^{n} \text{poly}(n) k \cdot 2^{n-1} = O^*(2^n).
\]

In this paper, we consider a family of problems that not only satisfy Property 1, but also the Dichotomic DPAS property below.

**Property 3 (Dichotomic DPAS).** Let \(t_0 \in T\). Problem \(P([n], t_0)\) can be solved by Dichotomic DPAS if there exist three functions \(t_1: 2^n \times 2^n \times T \rightarrow T\), \(t_2: 2^n \times 2^n \times T \rightarrow T\) and \(g: 2^n \times 2^n \times T \rightarrow \mathbb{R}\), computable in polynomial time, such that, for all \(J \subseteq [n]\) of even cardinality:

\[
\text{OPT}[J, t_0] = \min_{X \subseteq J} \left\{ \text{OPT}[X, t_1(J, X, t_0)] + g(J, X, t_0) + \text{OPT}[J \setminus X, t_2(J, X, t_0)] \right\}
\]

**Proof.** We compute Equation (3) for all \(J\) such that \(|J| = 2^k\), and for all \(t \in T\), starting from \(k = 1\) to \(k = \log_2(n)\). For a given \(J\), the values \(\{\text{OPT}[X, t'] : X \subseteq J, |X| = |J|/2, t' \in T\}\) are known, so \(\text{OPT}[J, t]\) is computed in time \(\text{poly}(n) 2^{2k}\) according to Equation (2) (the computation of \(t_1, t_2\) and \(g\) is polynomial). Thus, computing all \(\text{OPT}[J, t]\) for any \(J\) of size \(2^k\) and \(t \in T\) is done in
time \(|T|\text{poly}(n)\binom{2k}{n} \binom{n}{2k}\). Eventually, the total complexity is equal to

\[ B(n) = \frac{|T|\text{poly}(n)}{B(n)} \sum_{k=1}^{\log_2(n)} \left( \frac{2^k}{n} \right)^{n/2} \] .

We show that \(2^n B(n) \to 0\). Let us first consider the sub-sequence \((B(2^i))_{i \in \mathbb{N}}\). For \(n = 2^i\), a lower bound of \(B(n)\) is the sum of the two last terms: \(B(n) > |T|\text{poly}(n)\binom{n}{n/2} \approx C|T|\text{poly}(n) \binom{2^n}{n/2}\), where \(C\) is a constant. Moreover, the sequence \((B(n))_{n \in \mathbb{N}}\) is increasing. Thus, \(C\) dominates \(n \to 2^n\) asymptotically, namely \(C(n) = \omega(2^n)\). \(\square\)

Notice that solving \(\mathcal{P}\) using only Dichotomic DPAS is worse than using only DPAS. However, we describe in the next section a hybrid algorithm we call Quantum Dichotomic DPAS (Q-DDPAS) that improves the complexity of solving \(\mathcal{P}\) by combining DPAS and Dichotomic DPAS with Grover search. Before introducing this algorithm, we give some examples of single-machine scheduling problems that satisfy the Dichotomic DPAS property and then can be solved with Q-DDPAS.

### 2.2 Scheduling Examples

Let us begin with the scheduling problem with deadline constraints and minimization of the total weighted completion time.

**Example 5 (Minimizing the total weighted completion time with deadlines).** For each job \(j \in [n]\), we are given a weight \(w_j\), a processing time \(p_j\), and a deadline \(d_j\). We note \(p(J) = \sum_{j \in J} p_j\) and \(T = \sum_{j \in J} p_j\). For each job \(j \in [n]\) and \(t \in T\), we consider the problem \(P(J, t)\) where

\[ \Pi(J, t) = \left\{ \pi \in \mathcal{S}_J | C_j(\pi) \leq d_j - t, \forall j \in J \right\}, \]

where \(C_j\) is the completion time of job \(j\), and for \(\pi \in \Pi(J, t)\):

\[ f(\pi, J, t) = \sum_{j \in J} w_j(C_j(\pi) + t) . \]

\(P(J, t)\) represents the problem of finding the best feasible solution for jobs in \(J\) supposing that starting time is \(t\), and not 0 as usual. Our problem of interest is \(\mathcal{P} = \mathcal{P}([n], 0)\), often referred to as \(\mathcal{P} = [n], 0\). This family of problems also satisfies the Dichotomic DPAS. Indeed, Equation (2) is valid with:

\[ \forall j \subseteq [n], \forall t, \forall \pi \in \Pi(J, t) \]

\[ h(J, j, t) = \left\{ \begin{array}{ll} w_j(p(j) + t) & \text{if } d_j \geq p(j) + t \\ + \infty & \text{else} \end{array} \right. \]

where the computation of \(h\) is polynomial (linear). This family of problems also satisfies the Dichotomic DPAS property. Indeed, Equation (3) is valid for the following functions:

\[ \forall J \subseteq [n], \forall t, \forall \pi \in \Pi(J, t) \]

\[ t_1(J, X, t) = t \]

\[ t_2(J, X, t) = t + p(X) \]

\[ g(J, X, t) = 0 \]

We present another problem that satisfies the Dichotomic DPAS property, which is the scheduling problem with minimization of the total weighted tardiness.

**Example 6 (Minimizing the total weighted tardiness).** For each job \(j \in [n]\), we are given a weight \(w_j\), a processing time \(p_j\), and a due date \(d_j\). We note \(p(J) = \sum_{j \in J} p_j\). Let \(T = \sum_{j \in J} p_j\). For each \(J \subseteq [n]\) and \(t \in T\), we consider the problem \(P(J, t)\) where

\[ \Pi(J, t) = \mathcal{S}_J, \]

and for \(\pi \in \Pi(J, t)\):

\[ f(\pi, J, t) = \sum_{j \in J} w_j \max(0, C_j(\pi) - d_j + t) , \]

where \(C_j\) is the completion time of job \(j\), and \(\max(0, C_j - d_j + t)\) represents the tardiness of job \(j\) for the effective due date \(d_j - t\). Our problem of interest is \(\mathcal{P} = \mathcal{P}([n], 0)\), often referred to as \(\mathcal{P} = [n], 0\). This family of problems can be solved by DPAS. Indeed, Equation (2) is valid with:

\[ \forall J \subseteq [n], \forall t, \forall \pi \in \Pi(J, t) \]

\[ h(J, j, t) = w_j \max(0, p(j) - d_j + t) , \]

where the computation of \(h\) is polynomial (linear). This family of problems also satisfies the Dichotomic DPAS. Indeed, Equation (3) is valid for the following functions:

\[ \forall X \subseteq J \subseteq [n], \forall t, \forall \pi \in \Pi(J, t) \]

\[ t_1(J, X, t) = t \]

\[ t_2(J, X, t) = t + p(X) \]

\[ g(J, X, t) = 0 \]

We end with the example of the scheduling problem with precedence constraints and minimization of the total weighted completion time.

**Example 7 (Minimizing the total weighted completion time with precedence constraints).** We are given, for each job \(j \in [n]\) a processing time \(p_j\) and a weight \(w_j\), and a set of precedence constraints \(E = \{(i, j) : i < j\}\) that contains all pairs of jobs \((i, j)\) such that \(i\) precedes \(j\). We note \(p(J) = \sum_{j \in J} p_j\). Let \(T = \{0\}\). Here, the family of problems under consideration is indexed only by the chosen subset of \([n]\). Thus, for each \(J \subseteq [n]\), we consider the problem \(P(J, 0)\) where

\[ \Pi(J, 0) = \left\{ \pi \in \mathcal{S}_J | \pi \text{ respects } E \right\} , \]

and for \(\pi \in \Pi(J, 0)\):

\[ f(\pi, J, 0) = \sum_{j \in J} w_j C_j(\pi) , \]

where \(C_j\) is the completion time of job \(j\). Our problem of interest is \(\mathcal{P} = \mathcal{P}([n], 0)\), often referred to as \(\mathcal{P} = [n], 0\). This family of problems also satisfies the Dichotomic DPAS. Indeed, Equation (3) is valid for the following functions:

\[ \forall X \subseteq J \subseteq [n] \]

\[ t_1(J, X, 0) = 0 \]

\[ t_2(J, X, 0) = 0 \]

\[ g(J, X, 0) = \left\{ \begin{array}{ll} + \infty & \text{if } \exists (j, k) \in E \mid k \in J \\ p_j(p(J)) & \text{else} \end{array} \right. \]

where the computation of \(h\) is polynomial (quadratic). This family of problems also satisfies the Dichotomic DPAS. Indeed, Equation (3) is valid for the following functions:

\[ \forall X \subseteq J \subseteq [n] \]

\[ t_1(J, X, 0) = 0 \]

\[ t_2(J, X, 0) = 0 \]

\[ g(J, X, 0) = \left\{ \begin{array}{ll} + \infty & \text{if } \exists (j, k) \in E \mid j \in J \setminus X \text{ and } k \in X \\ p_j(p(J)) & \text{else} \end{array} \right. \]
where the computation of $g$ is polynomial (quadratic).

3 QUANTUM DICHOTOMIC DPAS ALGORITHM

In this section, we introduce a hybrid bounded-error algorithm called Quantum Dichotomic DPAS (Q-DDPAS) that solves scheduling problems satisfying the Dichotomic DPAS property described in the last section. It is an adaptation of the algorithm in [2]. We describe the quantum part of our algorithm in the gate-based quantum computing model, namely, as a quantum circuit decomposed into single and two-qubit quantum gates. The computational time of such a quantum circuit is quantified by the number of these elementary quantum operations [12]. Henceforth, we assume to have random access to quantum memory (QRAM) [7]. Notice that this is a strong assumption because QRAM is not available on current universal quantum hardware and is not expected to be so in the near future.

3.1 Preliminaries

We begin with some notions of complexity for quantum circuits and some notations for the description of Q-DDPAS.

**Definition 8.** Let us consider a family of quantum circuits $(Q_n)_{n \in \mathbb{N}}$ of complexity $O(C(n))$, meaning that $Q_n$ is a circuit that applies on $n$ qubits and contains $f(n)$ universal quantum gates, where $f(n) = O(C(n))$. This family is efficient if $C(n) = n^a$ for $a > 0$.

**Observation 9 (Complexity of quantum circuits).** Let $U_1$ and $U_2$ be two quantum circuits, with complexity $O(C_1(n))$ and $O(C_2(n))$, respectively. The complexity of the composition $U_1 \cdot U_2$ is

$$O(C_1(n) + C_2(n)) = O(\max(C_1(n), C_2(n))).$$

The tensor product $U_1 \otimes U_2$ has the same complexity.

**Observation 10 (Classical algorithm into quantum circuit).** Any classical algorithm $\mathcal{A}$ can be described as a quantum circuit $U_{\mathcal{A}}$. The complexity of $U_{\mathcal{A}}$ is equal to the complexity of $\mathcal{A}$.

We define two useful sets for the description of our algorithm, both indexed by a subset and a parameter, $(J, t)$. Essentially, the first set $\Lambda(J, t)$ contains all the possible balanced bi-partitions of $J$ and the associated parameter values of $t_1$ and $t_2$. The second set $\Omega(J, t)$ contains the optimal solutions for each bi-partition in $\Lambda(J, t)$.

**Definition 11 (Sets $\Omega$ and $\Lambda$).** For $J \subseteq [n]$ such that $|J|$ is even and for $t \in T$, we define the set

$$\Lambda(J, t) = \left\{ (X, t_1(J, X, t), J \setminus X, t_2(J, X, t)) : X \subseteq J, |X| = \frac{|J|}{2} \right\},$$

and the set

$$\Omega(J, t) = \left\{ (X, \text{OPT}(X, t_1(J, X, t), J \setminus X, t_2(J, X, t), t)) : X \subseteq J, |X| = \frac{|J|}{2} \right\}.$$

Let us introduce the quantum circuits that constitute the building blocks of our algorithm, and let us provide for each of them their complexity. The two first circuits $U_{\Lambda}$ and $U_{\Omega}$ amount to put into uniform superposition the elements of $\Lambda$, respectively $\Omega$.

**Definition 12 (Circuit $U_{\Lambda}$).** For $J \subseteq [n]$ such that $|J|$ is even, and for $t \in T$, we define $U_{\Lambda}$ as follows:

$$U_{\Lambda}(J) |t \rangle |0\rangle = |J\rangle |t\rangle \sum_{(J, t_1, t_2) \in \Lambda(J, t)} \frac{1}{\sqrt{|\Lambda(J, t)|}} |t_1\rangle |t_2\rangle |0\rangle.$$

Notice that we index the objects that represent sets by $s$, and the objects that represent parameters in $T$ by $t$.

**Property 13 (Complexity of $U_{\Lambda}$).** The complexity of $U_{\Lambda}$ is polynomial in the size of the input.

**Proof.** First, let us prove that the construction of the superposition of subsets of $J$ of size $|J|/2$ is polynomial. Let $J \subseteq [n]$ of size $m$ (we suppose $m$ to be even). Let us prove that the construction of a quantum superposition of balanced bi-partition (both subsets are $\frac{m}{2}$) of $J$ can be done in polynomial time.

We know that there are $\binom{m}{m/2}$ balanced bi-partitions. It is possible to construct implicitly a bijection

$$\sigma : [1, \ldots, \binom{m}{m/2}] \rightarrow \left\{ (X, J \setminus X) : |X| = \frac{m}{2} \right\},$$

that enumerates the balanced bi-partitions (it requires making a bijection between $J$ and $[1, \ldots, m]$). Notice that the construction of this bijection does not depend on the values of $J$. For a given $i$ in $[1, \ldots, \binom{m}{m/2}]$, the computation of $\sigma(i) = (X_i, J \setminus X_i)$ is polynomial. Thus,

$$|i\rangle |0\rangle \rightarrow |i\rangle |X_i\rangle |J \setminus X_i\rangle$$

is done with a quantum circuit with a polynomial complexity, for any $i$. It implies that

$$\sum_{i=1}^{\binom{m}{m/2}} |i\rangle |0\rangle \rightarrow \sum_{i=1}^{\binom{m}{m/2}} |i\rangle |X_i\rangle |J \setminus X_i\rangle$$

is also done in polynomial time, resulting from the application of Hadamard gates.

Eventually, the computation of functions $t_1$ and $t_2$ is polynomial (see hypothesis in Dichotomic DPAS Property 3). Thus, the complexity of $U_{\Omega}$ is polynomial.

**Definition 14 (Circuit $U_{\Omega}$).** For $J \subseteq [n]$ such that $|J|$ is even, and for $t \in T$, we define $U_{\Omega}$ as follows:

$$U_{\Omega}(J) |t \rangle |0\rangle = |J\rangle |t\rangle \sum_{\omega \in \Omega(J, t)} \frac{1}{\sqrt{|\Omega(J, t)|}} |\omega\rangle.$$

**Property 15 (Complexity of $U_{\Omega}$).** Let $J$ be the input set. If we suppose to have stored in the QRAM the values $\text{OPT}(X, t)$ for all $X \subseteq J$ such that $|X| = n/4$ and for all $t \in T$, the complexity of $U_{\Omega}$ is polynomial in the size of the input.

**Proof.** The proof follows essentially the same lines as the proof of Property 13. The quantum superposition of subsets is done in polynomial time, and instead of computing $t_1$ and $t_2$, we get values in the QRAM in constant time.
\[ r(\omega_1, \omega_2, \omega_3, \omega_4) = \alpha_1^2 + \omega_2^2 + g(\omega_1 \cup \omega_2, \omega_3, \omega_4). \]

We note \( U_r \), the quantum circuit corresponding to \( r \), namely:
\[ V(\omega_1, \omega_2, \omega_3, \omega_4) \in \Omega(J, t), U_r |0\rangle = |0\rangle |r(\omega)|, \]
where \( |0\rangle = |\omega_1\rangle |\omega_2\rangle |\omega_3\rangle |\omega_4\rangle \) is encoded in five registers. Notice that we index the objects that represent numerical values by \( v \).

**Property 17 (Complexity of \( U_r \)).** The complexity of \( U_r \) is polynomial in the size of the input.

**Proof.** The computation of \( g \) is polynomial (see hypothesis in Dichotomic DPAS Property 3). It implies that the hypothesis of \( r \) is polynomial, and thus that \( U_r \) has a polynomial complexity (Property 10). \( \square \)

**Definition 16 (Circuit \( U_{QMF} \)).** Let \( f : [n] \rightarrow \mathbb{Z} \) be a function and let \( U_f \) be its corresponding quantum circuit, specifically,
\[ U_f |i\rangle |0\rangle = |i\rangle |f(i)|, \forall i \in [n]. \]
We note \( U_{QMF}[U_f] \) the quantum circuit corresponding to the Quantum Minimum Finding algorithm [5] thatcomputes with high probability the minimum value of \( f \) and the corresponding minimizer:
\[ U_{QMF}[U_f] \sum_{i=1}^{n} \frac{1}{\sqrt{n}} |i\rangle |0\rangle = \sum_{i=1}^{n} \frac{1}{\sqrt{n}} |i\rangle \arg\min_{j \in [n]} \{f(i)\} \min_{j \in [n]} \{f(i)\}. \]

**Property 19 (Complexity of \( U_{QMF} \)).** The complexity of the Quantum Minimum Finding algorithm is \( O(\sqrt{n} \cdot C_f(n)) \), where \( n \) is the size of the domain of \( f \) and \( O(C_f(n)) \) is the complexity of the circuit \( U_f \). Thus, the complexity of \( U_{QMF}[U_f] \) is
\[ O(\sqrt{n} \cdot C_f(n)) \]
according to Observation 10.

In the next section, we use some notations as follows. Let \( \mathcal{R} = \{|q_1\}, \ldots, |q_n\rangle \) be a register of \( n \) qubits and \( U \) be an operator acting on \( k \) qubits, with \( k < n \). Let \( I \) be a \( k \)-tuple of distinct indices in \([n], I = (i_1, \ldots, i_k)\). We denote by \( U^I \) the operator acting on the full register \( \mathcal{R} \), that applies \( U \) on \( |q_{i_1}\rangle \ldots |q_{i_k}\rangle \), and applies \( I_d \) on the remaining qubits. For instance, if \( I \) is the tuple of contiguous indices \((3, \ldots, k + 3)\) with \( k < n - 3 \), then

\[ U^I = I_d^{k+2} \otimes U \otimes I_d^{n-k-3}. \]

For \( I = (i_1, \ldots, i_k) \) and \( J = (j_1, \ldots, j_l) \) two distinct tuples in \([n] \) (\( k \)-tuple and \( l \)-tuple where \( i \neq j \)), \( \forall (i, j) \in I \times J \), we note \( I \otimes J \) the concatenation of \( I \) and \( J \), namely \( I \otimes J = (i_1, \ldots, i_k, j_1, \ldots, j_l) \).

Let us denote the indexes related to the quantum circuit \( U_f \) as
\[ U_f |i\rangle |0\rangle = |i\rangle |f(i)\rangle. \]

To clarify the computations detailed next, we index the corresponding QMF operator as \( U_{QMF}[U_f^I] \). We omit the index \( J \) because this is an auxiliary register that does not appear in the output of \( U_{QMF}[U_f] \).

### 3.2 Description of the Algorithm

We describe the Quantum Dichotomic DPAS (Q-DDPAS) algorithm as an adaptation of [2] for scheduling problems satisfying the Dichotomic DPAS property. Without loss of generality, we assume that 4 divides \( n \). The hybrid quantum-classical algorithm Q-DDPAS consists of two steps:

1. **Classical part:** For each \( X \) of size \( n/4 \) and all \( t \in T \), solve by classical DPAS the problem \( P(X, t) \). Store the results in the QRAM as tuples \((X, t, OPT[X, t], \pi^*[X, t])\), where \( \pi^*[X, t] \) is the optimal permutation corresponding to \( OPT[X, t] \).
2. **Quantum part:**
   - Apply quantum circuit
   \[ U_{QDDPAS} = U_{recur} \cdot U_{ini} \]
   to the initial state
   \[ |\text{ini}\rangle = |[n]\rangle |0\rangle |0\rangle |0\rangle |0\rangle |0\rangle |0\rangle |0\rangle |0\rangle. \]
   where the tuples indexing the different registers are decomposed as follows:
   \[ t^1 = t^1 \otimes t^1 \]
   \[ t^2 = t^2 \otimes t^2 \]
   \[ t^3 = t^3 \otimes t^3 \]
   \[ t^4 = t^4 \otimes t^4 \]
   and where
   \[ U_{ini} = (U^T \otimes U^T) \cdot U_{\Lambda}, \quad (4) \]
   and
   \[ U_{recur} = U_{QMF}^r \]
   \[ |\text{recur}\rangle = |[n]\rangle |0\rangle. \]

(b) Measure register of indexes \( I_d^0 \) to find the optimal value \( OPT[n, 0] \).

The main idea of this algorithm is as follows. First, we compute classically by DPAS the optimal values of all subproblems scheduling with \( \frac{n}{4} \) jobs. Second, we call recursively two times QMF to find optimal values of subproblems scheduling with \( \frac{n}{2} \) jobs and eventually with \( n \) jobs (corresponding to the initial problem).

**Theorem 20.** The bounded-error Q-DDPAS algorithm solves \( \mathcal{P} \) in \( O^*(|T| \cdot 1.754^n) \).

The proof of Theorem 20 relies on the two lemmas introduced next. However, before stating and proving these lemmas, we observe that the complexity of Q-DDPAS can be further reduced by performing a third call to Dichotomic DPAS recurrence (3) as suggested in [2].

**Observation 21.** A slight modification of Q-DDPAS reduces the complexity to \( O^*(|T| \cdot 1.728^n) \).
For the sake of clarity, we will prove Observation 21 only after having proved Theorem 20. We now introduce the two lemmas necessary to prove Theorem 20.

Lemma 22. The optimal value of $P$ is stored in the register of indexes $I^n_2$ by Q-DDPAS with high probability.

Proof. We provide next a sketch of the proof, referring to Appendix B for the details of the computations. The main idea is to compute the first terms by classical DPAS, and then apply recursively twice Equation (3), which is solved by QMF:

- **Classical part:** Compute by classical DPAS the values $OPT[X,t]$ for all $X$ of size $n/4$ and for all $t \in T$. Store the results in the QRAM.
- **Quantum part:**
  - For each $J$ of size $n/2$ and $t \in T$, compute $OPT[J,t]$ through Equation (3) combined with QMF.
  - Compute $OPT[[n],0]$ with Equation (3) combined with QMF.

We now give some intuition on the effect of the quantum circuit $U_{QDDPAS} = U_{rec}U_{ini}$ and start by explaining the effect of $U_{ini}$ defined in (4). The application of $U_{ini}$ superposes all elements of $\Lambda([n],0)$ in the registers of indexes $I^2$ (partition $J$) and $I^4$ (partition $[n]\setminus J$). This essentially amounts to superpose all the $n_{n/2}^n$ bi-partitions of $[n]$ where each partition is of size $n/2$ (parameters $t$ included). Then, we apply $U_D$ on register of index $I^2$, resp. $I^4$. This superposes all elements of $\Omega(J,t)$ (for a $J$ of size $n/2$ and $t \in T$ previously described in registers of indexes $I^2$, resp. $I^4$). This essentially amounts to superpose all the $n_{n/2}^n$ bi-partitions of $[n]$ where each partition is of size $n/2$, parameters $t$ included, and the optimal value associated already stored in the QRAM.

Let us explain the effect of $U_{rec}$ defined in (5). The application of $U_{QMF}[U_r]$ on a register encoding $(J,t)$ and the superposition of elements of $\Omega(J,t)$ stores (with high probability) in an output register $OPT[J,t]$ according to the Dichotomic DPAS Property 3. Thus, $U_{QMF}[U_r]$ on register of index $I^2$, resp. $I^4$, superposes all $OPT[J,t]$ in $I^2$, resp. $I^4$. In other words, the circuit $U_{QMF}[U_r]$ that appears in $U_{rec}$ superposes (with high probability) all optimal values of Equation (3) for $J$ of size $n/2$. Now that the optimal values are known for sets of size $n/2$ (before, we only knew optimal values for sets of size $n/4$), we apply one more time $U_{QMF}[U_r]$ on these new registers: it outputs $OPT[[n],0]$ on the register of index $I^n_2$ with high probability.

Lemma 23. The complexity of Q-DDPAS is $O^*([T] \cdot 1.754^n)$.

Proof. Let us compute the complexity of this algorithm. First, we compute the complexity of the classical part. The proof of Lemma 2 shows that solving all $OPT[X,t]$ for all $X$ of size $n/4$ and for all $t \in T$ is done by DPAS in time

$$|T| poly(n) \sum_{k=1}^{n/4} k \binom{n}{k} = O^*\left(|T| \left(\binom{n}{n/4}\right)\right).$$

Thus, because $O^*(\binom{n}{n/4}) = O^*(2^{0.811n})$ (see Equation (6)), the complexity of the classical part is $O^*([T] \cdot 2^{0.811n})$.

Second, let us compute the complexity of the quantum part (using Property 10).

- The complexity of $U_{ini}$ is polynomial in $n$. Indeed, $U_A$ is polynomial in $n$ (Property 13). Moreover, $U_B$ is also polynomial in $n$: the classical part stored in the QRAM all $OPT[J,t]$ for $X$ of size $n/4$ and $t \in T$ (Property 15).
- The complexity of $U_{rec}$ is $O^*\left(\sqrt[n/2]{n/4}\right)$. Indeed, both terms $U_{QMF}[U_r]$ in $U_{rec}$ have a polynomial complexity for $U_r$ and find the minimum of functions with a domain of size $(n/2)^4$. Thus, each complexity of these two factors is $O^*\left(\sqrt[n/2]{n/4}\right)$, and so is the complexity of the tensor product. The circuit $U_{rec}$ has the same complexity because of the composition with $U_r$ that is polynomial. The circuit $U_{rec}$ finds the minimum of a function with a domain of size $(n/2)^4$ described by the corresponding quantum circuit $U_{rec}$ above. Thus, its complexity is $O^*\left(\sqrt[n]{n/2}\right)$.

Because $O^*\left(\sqrt[n]{n/2}\right) = O^*(2^{0.754n})$ (see Equation (8)), the complexity of the quantum part is $O^*\left(2^{0.754n}\right)$.

Eventually, the complexity of Q-DDPAS is

$$O^*(2^{0.754n} + |T| \cdot 2^{0.811n}) = O^*\left(|T| \cdot 2^{0.811n}\right) = O^*\left(|T| \cdot 1.754^n\right).$$

Proof of Lemma 23. Follows directly from Lemmas 22 and 23. □

Proof of Observation 21. The slight modification of Q-DDPAS amounts to add a level of recurrence in the quantum part, but instead of searching for the best concatenation among all the bi-partition of size $(n/8,n/8)$ (i.e. solving Equation (3) for $J = n/4$), we search for the best concatenation among all the bi-partition of size $(0.945 \cdot \frac{n}{8}, 0.055 \cdot \frac{n}{8})$, i.e. solving

$$OPT[J,t] = \min_{X \subseteq J \atop |X| = 0.34|J|} \left\{ OPT[X,t_1(J,X,t)] + g(J,X,t) + OPT[J \setminus X, t_2(J,X,t)] \right\}.$$ 

This further recurrence implies that:

- the classical part computes $OPT[J,t]$ for $J$ of size $0.945 \cdot \frac{n}{8}$ and $0.055 \cdot \frac{n}{8}$. Its complexity is then $O^*(|T|\leq 0.945 \cdot \frac{n}{8}) = O^*\left(|T| \cdot 1.789n\right)$ (see Equation (7)).
- the quantum part applies three levels of recurrence of QMF, finding the minimum over functions with a domain of size $(n/2^2) \cdot (n/4)$ and $(n/2^4)$ respectively. Its complexity is
then \( O^*(\sqrt{\frac{n}{n/2}} \sqrt{\frac{n/2}{n/4}}) = O^*(2.789n^2) \) (see Equation (9)).

The quantum part and the classical part have the same complexity, thus the total complexity of Q-DDPAS is the same, namely \( O^*(2.789n^2) = O^*(1.728^7) \).

We summarized in Table 1 the complexities of solving the scheduling problems studied in Section 2 with Q-DDPAS and compare them with the complexities of the best-known current classical algorithms. Q-DDPAS improves the complexity of the exponent but sometimes at the cost of a pseudo-polynomial factor (\( \sum p_j \) for problems \( 1\{d_j\} \sum_j w_j C_j \) and \( 1\{d_j\} \sum_j w_j T_j \)).

4 CONCLUSION

This paper extends the hybrid algorithm of [2] to scheduling problems that satisfy the Dichotomous DPAS property. Such problems, which are often solved in \( O^*(2^n) \) by classical DPAS, are solved by our bounded-error algorithm in \( O^*(\sqrt{T} \cdot 1.754^n) \), where \( T \) is meant to be at most pseudo-polynomial in the size of the problem. We detail the application of the resulting hybrid algorithm on three single-machine scheduling problems \( 1\{d_j\} \sum_j w_j C_j \) and \( 1\{d_j\} \sum_j w_j T_j \), showing a reduction of the exponent compared to the best-known classical complexity. We notice that a pseudo-polynomial factor appears in the complexity of two out of three problems. Future works will seek to extend these results to other scheduling problems for which the Dichotomous DPAS holds, such as the 3-machine flowshop scheduling problem [14].

REFERENCES


A NOTATIONS AND UPPER BOUNDS

In what follows, we use the notation

\[
\left(\begin{array}{c}
n \\
 k \\
 i \\
\end{array}\right) = \frac{n!}{k!i!}
\]

We also note the binary entropy of \( \epsilon \in [0, 1] \) the quantity \( H(\epsilon) = -\epsilon \log_2(\epsilon) + (1-\epsilon) \log_2(1-\epsilon) \). We give some useful upper bounds of binomial coefficients [2]:

\[
\forall k \in \left[ 1, \frac{n}{2} \right], \quad \left(\begin{array}{c}
n \\
 k \\
\end{array}\right) \leq 2^{H\left(\frac{k}{n}\right)} n,
\]

\[
\forall k \in \left[ \frac{n}{2}, n \right], \quad \left(\begin{array}{c}
n \\
 k \\
\end{array}\right) \leq 2^{H\left(\frac{k}{n}\right)} n.
\]

Observe that \( \left(\begin{array}{c}
n \\
 \frac{n}{4} \\
\end{array}\right) \) is bounded above by \( 2^{H\left(\frac{\frac{n}{4}}{n}\right)} n \), where

\[
H\left(\frac{\frac{n}{4}}{n}\right) = \frac{1}{4} \log_2\left(\frac{1}{4}\right) + \frac{3}{4} \log_2\left(\frac{3}{4}\right) = 0.811,
\]

so we obtain

\[
\left(\begin{array}{c}
n \\
 \frac{n}{4} \\
\end{array}\right) \leq 2^{0.811n} \leq 2^{0.789n}.
\]

In the same way, we can show that

\[
\left(\begin{array}{c}
n \\
 \frac{n}{2} \\
\end{array}\right) \leq 2^{0.789n}.
\]

Similarly, \( \left(\begin{array}{c}
n \\
 \frac{n}{2} \\
\end{array}\right) \) is bounded above by \( \sqrt{2^{H\left(\frac{\frac{n}{2}}{n}\right)}} n \), where

\[
H\left(\frac{\frac{n}{2}}{n}\right) = \frac{1}{2} H\left(\frac{1}{2}\right) + H\left(\frac{1}{2}\right) = 2^{\frac{1}{2} H\left(\frac{1}{2}\right)} n,
\]

so we obtain

\[
\left(\begin{array}{c}
n \\
 \frac{n}{2} \\
\end{array}\right) \leq 2^{0.75n}.
\]

In the same way, we can show that

\[
\left(\begin{array}{c}
n \\
 \frac{n}{4} \\
\end{array}\right) \leq 2^{0.789n}.
\]

B DETAILED PROOF OF LEMMA 22

Next, we compute \( U_{\text{rec}}^\dagger U_{\text{ini}}^\dagger \) and show that OPT \( [n], 0 \) is stored in register of indexes \( I^2 \). First, we compute \( U_{\text{ini}}^\dagger \).

\[
U_{\text{rec}}^\dagger U_{\text{ini}}^\dagger \mid \text{ini} \rangle = U_{\text{rec}}^\dagger \otimes \mathbb{I} \otimes \mathbb{I} \mid \text{ini} \rangle = \frac{1}{\sqrt{\lambda_1 \lambda_2}} \sum_{\lambda_1 \lambda_2 \in \Lambda([n],0)} \sqrt{\lambda_1 \lambda_2} \mid \lambda_1 \rangle \lambda_2 \rangle \mid 0 \rangle \mid 0 \rangle.
\]
Thus, we apply the second circuit of QMF.

\[
U_{\text{ini} | \text{ini}} = \left( U_{\text{QMF}}^{\text{ini}} \otimes U_{\text{QMF}}^{\text{ini}} \right) \cdot U_{\Lambda}^{\text{ini}} | \text{ini} \rangle
\]

\[
= \left( U_{\text{QMF}}^{\text{ini}} \otimes U_{\text{QMF}}^{\text{ini}} \right) \cdot \frac{1}{\sqrt{N}} \sum_{i \in \Omega} U_{\Lambda}^{\text{ini}} | i \rangle
\]

\[
= \frac{|i\rangle |0\rangle}{1} \sum_{i \in \Omega} \frac{1}{\sqrt{N}} U_{\Lambda}^{\text{ini}} | i \rangle
\]

Second, we apply the tensor product of the two first QMF to the previous state.

\[
\left( U_{\text{QMF}}^{\text{ini}} \otimes U_{\text{QMF}}^{\text{ini}} \right) \cdot U_{\Lambda}^{\text{ini}} | i \rangle
\]

\[
= \frac{1}{\sqrt{N}} \sum_{i \in \Omega} U_{\Lambda}^{\text{ini}} | i \rangle
\]

Thus, we apply the second circuit of QMF.

\[
U_{\text{rec2}} U_{\text{ini} | \text{ini}} = U_{\text{QMF}}^{\text{rec2}} \cdot U_{\Lambda}^{\text{ini}} | \text{ini} \rangle
\]

\[
= \frac{1}{\sqrt{N}} \sum_{i \in \Omega} \frac{1}{\sqrt{N}} U_{\Lambda}^{\text{ini}} | i \rangle
\]

According to definition of and the Dichotomic DPAS Property 3, the results stored in register of indexes \( t_3 \) is OPT[|n|, 0].

Notice that optimal permutation \( \pi' \) of \([n], 0\) can be rebuilt with registers of indexes \( t_3, t_2 \) and \( t_1 \) with the access to the results of the classical part in the QRAM.