

# Quantum Computationally Predicate-Binding Commitments with Application in Quantum Zero-Knowledge Arguments for **NP**

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## Abstract

A quantum bit commitment scheme is to realize bit (rather than qubit) commitment by exploiting quantum communication and quantum computation. In this work, we study the binding property of the quantum *string* commitment scheme obtained by composing a *generic* quantum *computationally-binding* bit commitment scheme *in parallel*. We show that the resulting scheme satisfies a stronger quantum computational binding property than the trivial honest-binding, which we will call *predicate-binding*. Intuitively and very roughly, the predicate-binding property guarantees that given any *inconsistent* predicate pair over a set of strings (i.e. no strings in this set can satisfy both predicates), if a (claimed) quantum commitment can be opened so that the revealed string satisfies one predicate with certainty, then the same commitment cannot be opened so that the revealed string satisfies the other predicate except for a negligible probability.

As an application, we plug a generic quantum perfectly (resp. statistically)-hiding computationally-binding bit commitment scheme in Blum's zero-knowledge protocol for the **NP**-complete language Hamiltonian Cycle. The quantum computational soundness of the resulting protocol will follow immediately from the quantum computational predicate-binding property of commitments. Combined with the perfect (resp. statistical) zero-knowledge property which can be similarly established as Watrous [Wat09], as well as known constructions of quantum computationally-binding bit commitment scheme, this gives rise to the first quantum perfect (resp. statistical) zero-knowledge *argument* system (with soundness error  $1/2$ ) for all **NP** languages merely relying on *quantum-secure one-way functions*.

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# 1 Introduction

Bit commitment is an important cryptographic primitive; it can be viewed as an electronic realization of a locked box [Gol01]. Roughly speaking, a bit commitment scheme has two stages, the commit stage and the reveal stage, providing two security guarantees: hiding and binding. Intuitively, the hiding property states that the commitments to 0 respective 1 are indistinguishable in the commit stage, whereas the binding property states that any (claimed) bit commitment cannot be opened (by the sender) as both 0 and 1 (except for a negligible probability) later in the reveal stage. Unfortunately, hiding and binding properties cannot be satisfied information-theoretically at the same time; one of them has to be *conditional*, e.g. based on complexity assumptions such as the existence of one-way functions.

Turning to the quantum setting, there are two *different* meanings of quantum bit commitment in the literature (depending on the context). The *first* refers to the *classical* realization of bit commitment that is secure against *quantum* attacks, or the post-quantum secure (classical) bit commitment [AC02, Unr16b, Unr16a]. The *second* refers to a realization of bit commitment by exploiting *quantum* features [BB84, BC90, DMS00, CLS01, CDMS04, KO09, KO11, CKR11, YWLQ15, FUYZ20, Yan20]; that is, now the honest parties are allowed to be quantum computers and exchange quantum messages. (But it is still a classical bit that we are trying to secure.) Clearly, the first meaning of quantum bit commitment can be viewed as a special case of the second one. In this paper, the term “quantum bit commitment” will be reserved for the second, more general meaning, which will also be the focus of this work.

The concept of quantum bit commitment is natural and sounds exciting. Though *unconditional* quantum bit commitment is still impossible [May97, LC98], as a compromise we can consider quantum bit commitment based on complexity assumptions like in the classical cryptography. Somewhat counter-intuitive at the first glance, but the binding property of a general quantum bit commitment scheme is inherently *weaker* than the classical binding property (that is guaranteed by a classical bit commitment scheme secure against classical attacks, which roughly states that any claimed bit commitment is bound to a unique bit that is typically referred to as the *committed value*). In more detail, this weakness of the quantum binding property comes from the possible superposition attack of the sender of quantum bit commitment, who may commit to an arbitrary *superposition* of bits 0 and 1, and later open the commitment as this superposition (rather than a classical 0 or 1) with certainty [DMS00, CDMS04]. By this kind of quantum superposition attack, a fixed quantum bit commitment is no longer bound to a unique classical bit any more. The quantum binding property that can be guaranteed by a general quantum bit commitment scheme is often referred to as *sum-binding* (named after [Unr16a]), which is equivalent to honest-binding w.r.t. the quantum bit commitment scheme of a *generic* form. (Refer to Definition 4 for a formal definition of a generic computationally-binding quantum bit commitment scheme.)

**Difficulties in basing security on quantum binding.** It is natural to ask what happen if we replace classical bit commitment with quantum bit commitment in cryptographic applications. Due to the weakness of the general quantum binding property as aforementioned, the security based on the classical binding property may deteriorate after the replacement.

In greater detail, note that in applications we typically commit to a binary string by committing it in a *bitwise* fashion; later, a *subset* of bit commitments may be opened for some verification. For example, it is helpful to keep GMW-type zero-knowledge protocols [Blu86, GMW91] in one’s mind. When quantum bit commitment is used, we can no longer say that a claimed quantum commitment to an  $m$ -bit string is bound to some  $m$ -bit string; instead, the committed value of such a quantum string commitment could be a superposition of a bunch of  $m$ -bit strings of the

form  $\sum_{s \in \{0,1\}^m} \alpha_s |s\rangle$ . One may tend to argue in security analysis that this superposition behaves similar to its induced probability distribution  $(|\alpha_s|^2)_{s \in \{0,1\}^m}$ : if this is true, then the classical security analysis can extend to the quantum setting straightforwardly. Unfortunately, this argument is not necessarily true, because a superposition is generally *not* equivalent to its induced probability distribution; in fact, this is usually where the quantum advantage comes from in algorithm design. If one goes into detail of the security analysis, one will find that a malicious quantum sender of commitments may attack by making the opening information (which is entangled with quantum commitments and their decommitments) about *which* bit commitments are to open as *what* value in an arbitrary superposition. By tuning this superposition, the sender may adjust the receiver’s acceptance probabilities in different verifications. This kind of superposition attack will make the security analysis based on the quantum binding property much harder than that based on the classical binding property (if possible).

**Why quantum bit commitment is interesting?** Besides the weakness as well as technical difficulties in security analysis mentioned above, another shortcoming of quantum bit commitment is that by today’s quantum technology, the physical realization of a general quantum bit commitment scheme is still far beyond our reach. In spite of this, quantum bit commitment still interests us for several reasons. First, since as early as 2000 researchers have come to realize that merely based on quantum-secure one-way functions/permutations, one can construct *non-interactive* quantum bit commitments of both flavors (i.e. statistical binding and statistical hiding), whose commit stage consists of just a single *quantum* message from the sender to the receiver [DMS00, KO09, KO11, Yan20]. It turns out that these constructions are not coincidences: recently, Yan [Yan20] has shown that any (interactive) quantum bit commitment scheme can be converted into a non-interactive one of a *generic* form. This is in contrast to the constant [MP12] or even polynomial [HHR07] number of rounds in the commit stage by classical constructions of bit commitment. Thus, using quantum bit commitment instead of the classical one in applications can potentially reduce the number of rounds of the interaction<sup>1</sup> while keeping the complexity assumption to the minimum.

More interestingly, Fang, Unruh, Yan and Zhou [FUZZ20] and Yan [Yan20] also observe that the (either statistical or computational) binding property of a generic non-interactive quantum bit commitment scheme is automatically *information-theoretically strict*. Here, the strictness of the quantum binding property extends the one in [Unr12] for a classical construction of bit commitment scheme in a straightforward way, which roughly states that not only the revealed value but also the *decommitment state* used in opening a quantum bit commitment are “unique”. We highlight that this quantum strictness originates from the *entanglement* between the commitment and its decommitment, as opposed to the *classical correlation* in the definition of the classical strict-binding [Unr12]. This also explains why even the quantum computational binding property could also be informal-theoretically strict<sup>2</sup>. It turns out that this strictness of the quantum binding property can play an important role in applications; in particular, it can help circumvent existing barriers only known for classical constructions, as confirmed in [FUZZ20]. (Jumping ahead, this paper will give yet another such an example; refer to Theorem 1.)

Overall, if we are optimistic about the development of quantum technology and believe that general quantum computation and communication will be available in future, then *the application of quantum bit commitment as a primitive in quantum cryptography* is worthy of study.

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<sup>1</sup>The round complexity of any cryptographic task might be one of the most important parameters.

<sup>2</sup>We note that the computational binding property of a classical bit commitment scheme cannot be information-theoretically strict: though it may be computationally hard to find an alternative opening, there actually *exist* a bunch of them!

**Progress and perspective towards basing security on quantum binding.** In the past two decades, there were only few works studying the security based on the binding property of a *general* quantum bit commitment scheme [YWLQ15]. Recently, some *generic* techniques to cope with the quantum *perfect/statistical* binding property are developed in [FUYZ20], by which in many cases the security based on the *classical* statistical binding property can be lifted to the quantum setting. Unfortunately, when it comes to the setting of quantum *computational* binding, the security remained elusive. To the best of our knowledge, we are aware of no such results before. In our opinion, the perhaps most important open question towards using quantum bit commitment as a primitive in quantum cryptography is:

*Can we base quantum security on the general computational binding property of quantum bit commitment?*

Based on the state-of-the-art knowledge, the answer to the question above is unclear. On one hand, it will be true if we can view the *superposition* of strings underlying quantum bit commitments as its induced *probability distribution* as aforementioned. Actually, this straightforward strategy indeed works (in many cases) when perfectly/statistically-binding quantum bit commitment is used [FUYZ20]. More positive evidences come from the success in various security analysis in the quantum random oracle model, in which adversaries can query a random oracle in an arbitrary superposition [BDF<sup>+</sup>11].

On the other hand, however, after a first attempt towards the security analysis, it turns out that for a naive analysis (r.f. subsection 1.2) to work it requires that the binding error be *sub-exponentially* or even *exponentially* small, rather than *negligibly* small as typical in cryptography. We will refer to this technical difficulty as “exponential curse”, which arises from that polynomial number of qubits could be in a superposition of exponentially many basis states. Moreover, the impossibility of general quantum rewindings [vdG97], as well as other related impossibility results w.r.t. classical constructions of bit commitment secure against quantum attacks [ARU14], may suggest a *negative* answer to the open question above.

One motivation of this work is to explore the application of quantum computationally-binding bit commitment in cryptographic applications, notably in constructing quantum zero-knowledge arguments for **NP** languages.

## 1.1 Our contribution

In spite of the technical difficulty and negative evidences just mentioned, we make some progress towards answering the main open question *affirmatively* in this work. Interestingly, our security analysis turns out to be completely *different* from the straightforward strategy of viewing the superposition of strings underlying quantum bit commitments as its induced probability distribution.

Specifically, our contribution is two-fold.

### 1. A *quantum* construction of perfect/statistical zero-knowledge *argument* system (with soundness error $1/2$ ) for all **NP** languages

We prove the following main theorem of this paper:

**Theorem 1** *Plugging a generic quantum perfectly (resp. statistically)-hiding computationally-binding bit commitment scheme in Blum’s protocol [Blu86] gives rise to a three-round public-coin quantum perfect (resp. statistical) zero-knowledge argument system for the **NP**-complete language Hamiltonian Cycle, with perfect completeness and soundness error  $1/2$ .*

Following [KO09, KO11, Yan20], since a generic quantum (statistically-hiding) computationally-binding bit commitment scheme can be constructed from quantum-secure one-way functions, the theorem above gives the *first* quantum statistical zero-knowledge argument for all **NP** languages based on the same assumption.

**Corollary 1** *If quantum-secure one-way function exists, then all languages in **NP** have a three-round public-coin quantum statistical zero-knowledge argument system with perfect completeness and soundness error  $1/2$ .*

Compared with classical GMW-type statistical zero-knowledge arguments (secure against classical attacks) for **NP**, our *quantum* construction reduces the rounds of the interaction *from polynomial to three*, thanks to the *non-iterativeness* of a generic quantum computationally-binding bit commitment scheme. Compared with the classical statistical zero-knowledge argument for **NP** secure against quantum attacks in [Unr16b, Unr16a], our quantum construction relies on a weaker (perhaps minimum) complexity assumption.

We highlight that our proof of Theorem 1 relies heavily on (though implicitly) that the (computational) binding property of a generic quantum bit commitment scheme is information-theoretically *strict*, as aforementioned. It is this strict-binding property that enables a simple quantum rewinding [YWLQ15, FUYZ20] to work even in our quantum *computational* soundness analysis, thus circumventing a barrier which is only known for classical constructions [ARU14].

## 2. A non-trivial computational binding property of the quantum *string* commitment scheme obtained by composing a generic quantum bit commitment scheme in parallel

A natural way to construct a string commitment scheme is to compose a bit commitment scheme in *parallel*, i.e. committing a string in a bitwise fashion. For the purpose of proving Theorem 1, we introduce a new binding property of quantum *string* commitments which we call “predicate-binding”. And we show that the parallel composition of a generic quantum computationally-binding bit commitment scheme gives rise to a quantum computationally predicate-binding string commitment scheme. It turns out that the quantum computational soundness of Blum’s protocol instantiated with a generic quantum computationally-binding bit commitment scheme (which is required towards establishing Theorem 1) can be easily based on the predicate-binding property of quantum string commitments.

In more detail, we first formalize a kind of predicates which we will call “pattern-predicates” (Definition 7): informally speaking, for a string to satisfy a pattern-predicate, it should exhibit a certain “pattern” somewhere. The *intuition* underlying our definition is that in typical applications of bit commitments, the receiver (of commitments) will check whether the value of the opened commitments will cause it to accept. For example, in Blum’s protocol the (honest) verifier’s verification corresponding to each challenge naturally induces a pattern-predicate.

With our definition of pattern-predicate, the *predicate-binding* property (Definition 8, or formally Definition 9) guarantees that given an arbitrary pair of *inconsistent* pattern-predicates on a set of strings of the same length (i.e. no strings in this set can satisfy both predicates), if a (claimed) quantum commitment can be opened such that the revealed string<sup>3</sup> satisfies one predicate with *certainty*, then the same commitment cannot be opened so as to satisfy the other predicate (except for a negligible probability)<sup>4</sup>.

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<sup>3</sup>Generally, the revealed value of a quantum string commitment could be a probability distribution over this set of strings.

<sup>4</sup>We note that the parallel composition of *classical* bit commitments secure against classical attacks gives a string

The proof of predicate-binding is the main technical contribution of this work, which is highly non-trivial; in particular, the trivial reduction (via a simple hybrid argument) from string binding to bit binding in the classical setting will fail completely here. Actually, for a technical reason we did *not* prove the *full* predicate-binding property (i.e. w.r.t. the most general inconsistent pattern-predicate pairs) in this work; rather, we can only show predicate-binding such that one predicate is allowed to be of the general form, whereas the other is subject to the restriction that it only depends on a *fixed* portion of the string (Theorem 2, or formally Theorem 3). In spite of this restriction, the predicate-binding property we obtain is sufficient for establishing Theorem 1. Any extension of our result is left as an open problem. We believe that quantum predicate-binding string commitments could be of independent interest and will be found useful elsewhere.

**A comparison with existing quantum computational string binding properties.** The parallel composition of a generic quantum bit commitment scheme trivially gives a quantum *honest-binding* string commitment scheme [YWLQ15]. Roughly speaking, the honest-binding property states that the honest commitment to a string cannot be opened as any other string (except for a negligible probability). Unfortunately, this binding property seems too weak to be useful in applications.

In [CDMS04], a so-called computational *f-binding* property w.r.t. a function  $f : \{0, 1\}^m \rightarrow \{0, 1\}^l$  for quantum string commitments is proposed, where integers  $l \leq m$ . Unfortunately, no constructions for quantum *f-binding* commitments are provided in [CDMS04]. Interestingly, we note that the predicate-binding property we established here implies the *f-binding* w.r.t. to any efficiently computable function  $f$  whose image is just the set  $\{0, 1\}$  (i.e.  $l = 1$ ), if we view preimages mapped to 0 as inducing one predicate while preimages mapped to 1 as inducing the other.

Damgård, Fehr and Salvail [DFS04] introduced the so-called *Q-binding* property for classical commitments secure against quantum attacks, which can be extended to quantum commitments in a straightforward way. Here, the “*Q*” stands for an arbitrary predicate whose form is close to our pattern-predicate<sup>5</sup>: very roughly, this predicate *Q* can be viewed as combining various pattern-predicates into one by introducing a “choice” parameter  $u$ , and the predicate-binding we establish here can also be viewed as the *Q-binding* w.r.t. the predicate *Q* of a special form such that  $|U| = 2$  and  $p_{\text{IDEAL}} = 1$ , by the notation used in [DFS04]. Constructions of *Q-binding* (classical) commitments in [DFS04] rely on stronger assumptions than quantum-secure one-way functions; in particular, one crucial assumption<sup>6</sup> on which it relies has a similar structure as the security game in defining *Q-binding*, which makes the security proof for *Q-binding* there much more straightforward than ours for predicate-binding here.

Unruh [Unr16b, Unr16a] introduced computational collapse-binding commitments secure against quantum attacks. However, a straightforward extension of collapse-binding to quantum commitments cannot hold generally, as discussed in more detail in [Yan20].

## 1.2 Technical overview

We sketch the soundness analysis of Blum’s protocol instantiated with a generic quantum computationally-binding bit commitment scheme, which is the key step towards establishing Theorem 1. Our goal

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commitment that is trivially predicate-binding secure against classical attacks. This is because the resulting string commitment (by the parallel composition) is bound to a unique classical string.

<sup>5</sup>As communicated by the authors of [DFS04] recently [DFS21], the definition of *Q-binding* in the conference version of [DFS04] has a flaw: it misses an additional information  $z$  as another input of the predicate *Q* to make it *efficiently computable*, and the sentence “We do not require *Q* to be efficiently computable” there should be removed.

<sup>6</sup>Namely, the third assumption in [DFS04, section “A General Framework”]

is to reduce the soundness of the resulting protocol to the predicate-binding property of quantum string commitment (Lemma 12).

We assume that readers are familiar with Blum’s protocol [Blu86], which is also sketched in subsection 2.3. In its soundness analysis, the (possibly cheating) prover’s first message constitutes a (claimed) quantum string commitment. The (honest) verifier’s acceptance conditions corresponding to the challenges 0 respective 1 induce two predicates on graphs with the same number of vertices as the input graph; when the input graph is not Hamiltonian, these two predicates will be *inconsistent* in the sense that no single graph can satisfy both of them simultaneously. Technically, at the heart of the reduction from the soundness of Blum’s protocol to the predicate-binding property of the quantum string commitment lies in a simple quantum rewinding technique that is similar to ones used in [Unr12, YWLQ15, FUYZ20], which basically states that if the verifier’s acceptance probability is high, then the quantum rewinding via reversible computations works. In greater detail, we can extend the quantum rewinding lemma that is suitable for the quantum statistical binding setting [FUYZ20] to the quantum computational binding setting (Lemma 2). We remark that though this extension is technically trivial, conceptually why it is possible relies heavily on that a generic quantum computationally-binding bit commitment scheme is *information-theoretical* strict-binding.

We are then left with showing that the parallel composition of a generic quantum computationally-binding *bit* commitment scheme indeed gives rise to a quantum computationally predicate-binding *string* commitment scheme (a special case in Lemma 10 and a more general case in Theorem 3). This is the main technical part of the paper. In the below, we first explain a technical difficulty towards this goal by a naive try, and then sketch at a high level how to overcome it. But before doing this, we first set up some notations that are necessary for our exposition.

**Notations.** A generic quantum bit commitment scheme can be represented by a quantum circuit pair<sup>7</sup>  $(Q_0, Q_1)$  performing on quantum registers  $(C, R)$ . To commit a bit  $b \in \{0, 1\}$ , in the commit stage the sender performs the quantum circuit  $Q_b$  on quantum registers  $(C, R)$  initialized in the state  $|0\rangle$ , and then sends the *commitment* register  $C$  to the receiver; later in the reveal stage, the sender sends the bit  $b$  together with the *decommitment* register  $R$  to the receiver, who then does the reversible computation (i.e. performing the quantum circuit  $Q_b^\dagger$ ) to decide whether to accept or not (i.e. checking whether the registers  $(C, R)$  return to the all  $|0\rangle$  state). Informally, we say that the quantum bit commitment scheme  $(Q_0, Q_1)$  is *computationally binding* if for any polynomial-time realizable unitary transformation  $U$  performing on the register  $R$ , the inner product  $|\langle 0|Q_1^\dagger U Q_0|0\rangle|$  is negligible; that is, unit vectors  $U Q_0|0\rangle$  and  $Q_1|0\rangle$  are almost orthogonal<sup>8</sup>.

To commit a string of length  $m$ , we commit it in a *bitwise* fashion using the scheme  $(Q_0, Q_1)$ . Let  $Q_s$  denote the corresponding quantum circuit used to commit the string  $s$ ; that is,  $Q_s = \bigotimes_{i=1}^m Q_{s_i}$ , which performs on  $m$  copies of the quantum registers  $(C, R)$ .

Let  $P_1, P_2$  be two (pattern-)predicates<sup>9</sup> on all  $m$ -bit strings. We use  $s \in P_1$  (resp.  $P_2$ ) to denote that the string  $s \in \{0, 1\}^m$  satisfies the predicate  $P_1$  (resp.  $P_2$ ). We say that two predicates  $P_1, P_2$  are *inconsistent* if no string  $s \in \{0, 1\}^m$  can satisfy both  $P_1$  and  $P_2$ . More details about the formalization of predicates are referred to subsection 4.1.

**A technical difficulty: exponential curse.** We first consider the *simplest* scenario, in which an

<sup>7</sup>For the moment, we drop the security parameter to simplify the notation.

<sup>8</sup>The formal definitions of a generic quantum bit commitment scheme and its computational binding property are referred to Definition 4. Here for simplification, we neglect the auxiliary input state that the cheating sender may receive.

<sup>9</sup>For the moment, we can think of them as efficiently computable predicates in the common sense for simplicity.



$m$ -bit string is firstly committed and later *all* (bit) commitments are to open. Note that a cheating sender can first prepare an arbitrary superposition of the form  $\sum_{s \in P_1} \alpha_s |s\rangle^D (Q_s |0\rangle)^{C^{\otimes m} R^{\otimes m}}$  (resp.  $\sum_{s \in P_2} \beta_s |s\rangle Q_s |0\rangle$ ) in registers  $(D, C^{\otimes m}, R^{\otimes m})$ , and then send all commitment registers  $C^{\otimes m}$  to the receiver in the commit stage. Later in the reveal stage, the sender sends the opening register  $D$ , together with all decommitment registers  $R^{\otimes m}$ , to the receiver. By this strategy, the sender can open all commitments as a *distribution* (which is determined by coefficients  $\alpha_s$ 's (resp.  $\beta_s$ 's)) of strings that satisfy the predicate  $P_1$  (resp.  $P_2$ ) with *certainty*. To show predicate-binding, we are sufficient to show that up to any *polynomial-time* realizable unitary transformation  $U$  that does not touch commitment registers  $C^{\otimes m}$  (which represents the sender's strategy in opening commitments), any two superpositions  $\sum_{s \in P_1} \alpha_s |s\rangle Q_s |0\rangle$  and  $\sum_{s \in P_2} \beta_s |s\rangle Q_s |0\rangle$  are almost *orthogonal*, i.e. their inner product is negligible, w.r.t. any inconsistent predicate pair  $(P_1, P_2)$ . A technical difficulty in showing this lies in that a potential exponential blow-up may occur in bounding this inner product. This difficulty is referred to as the *exponential curse* in [YWLQ15, FUYZZ20], which we believe is universal when one tries to base security on quantum binding; a similar difficulty also appears in [CDMS04]. Now let us go into some detail in the below.

By the computational binding property of the quantum bit commitment scheme  $(Q_0, Q_1)$ , the inner product  $|\langle 0 | Q_s^\dagger U Q_{s'} |0\rangle|$  where  $s \neq s'$  can be bounded by its binding error, which is negligible (as typical in cryptography). Thus, a naive way to bound the inner product

$$\left| \sum_{s \in P_1} \alpha_s^* \langle s | (\langle 0 | Q_s^\dagger U \sum_{s' \in P_2} \beta_{s'} |s'\rangle (Q_{s'} |0\rangle)) \right|$$

is first to expand it and bound each term indexed by  $(s, s')$  using the binding error bound (while neglecting its coefficient that can be bounded by 1), and then apply the triangle inequality. However, when there are *super-polynomial* (typically exponentially many) strings  $s \in P_1$  or  $s' \in P_2$ , this naive approach will fail.

Actually, whether the inner product above could really be bounded by some negligible quantity is questionable a priori. This is because generally, two superpositions of the form  $\sum_x \alpha_x |\phi_x\rangle$  and  $\sum_y \beta_y |\xi_y\rangle$ , where  $\{|\phi_x\rangle\}_x$  and  $\{|\xi_y\rangle\}_y$  are two orthonormal bases, are *not* necessarily almost orthogonal, even when  $|\phi_x\rangle$  and  $|\xi_y\rangle$  are almost orthogonal for each  $(x, y)$  pair. To see this, consider the following simple example. The Hilbert space is induced by  $m$  qubits, where  $\{|x\rangle\}_{x \in \{0,1\}^m}$  is the standard basis and  $\{H^{\otimes m} |y\rangle\}_{y \in \{0,1\}^m}$  is the Hadamard basis. Then consider an arbitrary vector in this space, which can be written as a superposition of basis vectors either in the standard basis or the Hadamard basis. Clearly, these two superpositions are actually the same vector, so that their inner product is one. But the inner product between  $|x\rangle$  and  $H^{\otimes m} |y\rangle$  for arbitrary  $x, y \in \{0,1\}^m$  is exponentially small! This example tells us that to bound the inner product aforementioned, we need to exploit the *structures* of the two superpositions (which are induced by the structures of predicates  $P_1$  and  $P_2$ ).

Similar technical difficulty also appears in the quantum statistical binding setting, where two generic techniques were invented to overcome this exponential curse: perturbation and hypothetical commitment measurement [YWLQ15, FUYZZ20]. Unfortunately, neither of them extend to the quantum computational binding setting studied here straightforwardly. For the reason, the *fundamental difference* between these two settings lies in that in the quantum statistical binding setting, the bit commitments to 0 respective 1 (stored in the commitment register  $C$ ) themselves are already *almost orthogonal*, and which will *never* be touched by the (possibly cheating) sender *after* they are sent. Thus, we can assume that commitments will *collapse* immediately by hypothetical commitment measurements at the moment they are sent; after the collapse, everything will be similar to that in the *classical* perfect binding setting. However, in the quantum computational

binding setting, commitments to 0 respective 1 could be *close or even identical*, where we are only guaranteed that in the reveal stage the *joint* states of the commitment register C and the decommitment register R are almost orthogonal. But the state of the decommitment register R can be affected by the sender’s operation *after* the commitment stage. In turn, the hypothetical-collapse trick to handle quantum statistically-binding commitments [FUYZ20] fails completely here.

In summary, new techniques are needed to establish the quantum computational predicate-binding property (if possible).

**Our approach.** Instead of considering the aforementioned inner product, now let us equivalently consider the *projection* of an arbitrary superposition of the form  $\sum_{s \in P_1} \alpha_s |s\rangle Q_s |0\rangle$ , up to any polynomial-time realizable unitary transformation  $U$  that does not touch commitment registers  $C^{\otimes m}$ , on the subspace induced by the predicate  $P_2$ , i.e.  $\sum_{s \in P_2} |s\rangle \langle s| \otimes (Q_s |0\rangle \langle 0| Q_s^\dagger)$ , which we also denote by  $P_2$ . Our goal then becomes to show that this projection is negligible. Our idea is based on the following *key observation*: when the predicate  $P_1$  is *sparse*, i.e. the number of the  $m$ -bit strings satisfying it is *polynomially* bounded, then combining a new *perturbation* technique (which looks similar but is inherently different from the one developed in the quantum statistical binding setting [YWLQ15, FUYZ20]) and the triangle inequality, we can bound the aforementioned projection by a negligible quantity. However, to remove this sparsity requirement, we still need to overcome the exponential curse. To this end, we need to take into account of the *coefficients* of the superposition, and make an essential use of the following *structure* of predicates  $P_1$  and  $P_2$ : to check whether a string satisfies  $P_1$  or  $P_2$ , *all* its bits are to examine.

For more technical details, we are to bound the norm

$$\left\| \sum_{s \in P_1} \alpha_s P_2 U (|s\rangle Q_s |0\rangle) \right\|,$$

where in the summation there could be exponentially many terms. At a high level, our *trick* is to order these terms properly in such a way that they can be treated as *leaves of a binary tree*, whose internal nodes will correspond to the summation of leaves of the subtree it determines; in particular, the root of the tree will correspond to the summation of all leaves, whose norm is just what we want to bound. We will actually bound the norm of all internal nodes, including the root, in a *bottom-up* fashion. It turns out that the accumulated error will grow only *linearly* in the *depth* of the tree, which is just  $m$ . The formal proof (of Lemma 10) is by induction on the depth of internal nodes.

As a final remark, we note that our security analysis did *not* achieve a totally *uniform* security reduction (from the quantum computational string predicate-binding property to the quantum computational binding property of a generic quantum bit commitment scheme); rather, we make an essential use of a certain amount of both classical and quantum *non-uniformity*. (More detail about this is referred to the discussion at the end of subsection 4.3.)

**Extension.** However, the (simplest) scenario (i.e. all commitments are to open) considered above is usually *not* sufficient for applications. This is because in many cases where bit commitments are used in a larger protocol, *not* all bit commitments are required to open for a verification. Even worse, positions of which bit commitments are to open may not even be fixed: they may depend on the party who plays the role of the (cheating) sender. For example, consider an execution of Blum’s protocol in which a Hamiltonian cycle is challenged to open.

Fortunately, we can extend the predicate-binding property established in the simplest case to a more general case in which it holds that for at least one predicate ( $P_1$  or  $P_2$ ), the positions of which bit commitments are to open for its verification are fixed, while the other predicate could

be arbitrary (Theorem 3). It turns out that this extension already suffices for our purpose of establishing Theorem 1.

For the formal proof of such extension, there are more technical issues we need to handle. (More detail is referred to the proof of Theorem 3 (in Appendix A)). Among others, we highlight that the computational binding property of a generic quantum bit commitment scheme needs to be strengthened (Lemma 5) for use.

*Organization.* We first give preliminaries in section 2. In section 3, we derive a more general computational binding property of a generic quantum bit commitment scheme from the standard one, which will be useful in the subsequent section 4, where we establish the computational predicate-binding property of the quantum string commitment scheme obtained by composing a generic quantum computationally-binding bit commitment scheme in parallel. As an application of the predicate-binding property, in section 5 we show that Blum’s zero-knowledge protocol for the NP-complete language Hamiltonian Cycle with a generic quantum computationally-binding bit commitment scheme plugged in is sound against any quantum computationally bounded prover. We conclude with section 6.

## 2 Preliminaries

A quantum system or register induces a Hilbert space. A quantum operation performing on a quantum system induces an operator acting on the Hilbert space associated with the system. In particular, a unitary operation induces a unitary transformation, and a binary projective measurement induces a projector (corresponding to the outcome one). We will *interchangeably* use quantum system and its induced Hilbert space, quantum operation and its induced operator. For example, we may say that a unitary transformation or a projector perform on or do not touch a quantum register.

**Notations.** We will explicitly write quantum register(s) as a *superscript* of an operator to indicate or highlight on which register(s) this operator performs. Similarly, we will also explicitly write quantum register(s) as a *superscript* of a quantum state to indicate or highlight in which register(s) this quantum state is stored. For example, let  $A$  be a quantum register. Then we may write  $U^A$ ,  $|\psi\rangle^A$  (resp.  $\rho^A$ ), to indicate that the operator  $U$  performs on the register  $A$ , the quantum pure (resp. mixed) state  $|\psi\rangle$  (resp.  $\rho$ ) is stored in the register  $A$ , respectively. We may also write  $U \otimes \mathbb{1}^A$  to highlight that the operation  $U$  does *not* touch the register  $A$ . But when it is clear from the context, we often drop such superscripts or the tensor product with the identity to simplify the notation; this in particular happens in many of derivations within our proofs, where we often write out registers as superscripts or the tensor product with the identity explicitly in the first step, while dropping them subsequently. When there are  $m$  copies of register  $A$ , and a unitary transformation  $U$  performs on the copies of the register  $A$  indexed by the subset  $T \subseteq \{1, 2, \dots, m\}$ , then we write  $\otimes T$  as the superscript, i.e.  $U^{A^{\otimes T}}$ ; in particular, when the subset  $T$  is the whole set, we then just write  $A^{\otimes m}$  to simplify the notation.

**Efficiently realizable quantum computation.** In this work, without loss of generality, we restrict to consider the following quantum computational model:

1. Quantum systems or registers are constituted of *qubits*.
2. There are only two kinds of quantum operations: *unitary* transformation and *projective* mea-

surement.

We also need to formalize *efficiently realizable* quantum operations. By [Yao93], any efficiently realizable quantum algorithm or unitary transformation can be formalized by a family of quantum circuits  $\{Q_n\}_{n \geq 1}$  such that:

1. Each gate of the quantum circuit  $Q_n$  comes from a pre-fixed finite, unitary, and universal quantum gate set, e.g. {Hadamard, phase, CNOT,  $\pi/8$ } [NC00].
2. Quantum circuit  $Q_n$  is of *polynomial* size (w.r.t. the index  $n$ ).
3. The quantum circuit family  $\{Q_n\}_{n \geq 1}$  can be uniformly generated, i.e. there exists a polynomial-time classical algorithm  $A$  which on input  $1^n$  outputs the description of the quantum circuit  $Q_n$ .

Since any *projective* measurement can be realized by first performing a unitary transformation, followed by a measurement of all qubits in the *standard* basis, we say that a projective measurement is *efficiently realizable* if the corresponding unitary transformation is efficiently realizable.

Any projector  $\Pi$  induces a binary measurement  $\{\Pi, \mathbb{1} - \Pi\}$ , which produces the outcome 1 (resp. 0) when the quantum state collapses into the subspace induced the projector  $\Pi$  (resp.  $\mathbb{1} - \Pi$ ). We say that the projector  $\Pi$  is *efficiently realizable* if its induced binary measurement is efficiently realizable.

**Quantum rewinding.** A quantum rewinding technique as stated in the lemma below is adapted from the one given in [FUYZ20] directly, whereas now we restrict to consider projectors and unitary transformations that are *efficiently realizable*. In spite of this, its proof follows the same line as the one in [FUYZ20].

**Lemma 2 (A quantum rewinding)** *Let  $\mathcal{X}$  and  $\mathcal{Y}$  be two Hilbert spaces. Unit vector  $|\psi\rangle \in \mathcal{X} \otimes \mathcal{Y}$ . Efficiently realizable projectors  $\Gamma_1, \dots, \Gamma_k$  perform on the space  $\mathcal{X} \otimes \mathcal{Y}$ , and efficiently realizable unitary transformations  $U_1, \dots, U_k$  perform on the space  $\mathcal{Y}$ . If  $1/k \cdot \sum_{i=1}^k \|\Gamma_i(U_i \otimes \mathbb{1}^X) |\psi\rangle\|^2 \geq 1 - \eta$ , where  $0 \leq \eta \leq 1$ , then*

$$\left\| (U_k^\dagger \otimes \mathbb{1}^X) \Gamma_k (U_k \otimes \mathbb{1}^X) \dots (U_1^\dagger \otimes \mathbb{1}^X) \Gamma_1 (U_1 \otimes \mathbb{1}^X) |\psi\rangle \right\| \geq 1 - \sqrt{k\eta}. \quad (1)$$

## 2.1 A generic quantum bit commitment scheme

We first need to define quantum (*in*)*distinguishability* based on the efficiently realizable quantum computation we fixed above. Our definition follows [Wat09].

**Definition 3 ((In)distinguishability of quantum state ensembles)** Two quantum state ensembles  $\{\rho_n\}_{n \geq 1}$  and  $\{\xi_n\}_{n \geq 1}$  are *quantum statistically* (resp. *computationally*) *indistinguishable* if for any quantum state ensemble  $\{\sigma_n\}_{n \geq 1}$  and any unbounded (resp. efficiently realizable) quantum algorithm  $D$  which outputs a single qubit that will be measured in the standard basis, it holds that

$$|\Pr[D(1^n, \rho_n \otimes \sigma_n) = 1] - \Pr[D(1^n, \xi_n \otimes \sigma_n) = 1]| < \text{negl}(n)$$

for sufficiently large  $n$ , where  $\text{negl}(\cdot)$  is some negligible function.

Following Yan [Yan20], the definition of a generic quantum computationally-binding bit commitment scheme is given as below.

**Definition 4 (A generic computationally-binding quantum bit commitment scheme)** A generic computationally-binding quantum bit commitment scheme is a two-party, two-stage protocol. It can be represented by an ensemble of polynomial-time uniformly generated quantum circuit pair  $\{(Q_0(n), Q_1(n))\}_{n \geq 1}$ . Specifically,

- The scheme involves two parties, a sender and a receiver, proceeding in two stages: a *commit* stage followed by a *reveal* stage.
- In the commit stage, to commit bit  $b \in \{0, 1\}$ , the sender performs the quantum circuit  $Q_b(n)$  on quantum registers (C, R) initialized in all  $|0\rangle$ 's state. Then the sender sends the *commitment register* C, whose state at this moment denoted by  $\rho_b(n)$ , to the receiver.
- In the (canonical) reveal stage, the sender announces  $b$ , and sends the *decommitment register* R to the receiver. The receiver then performs  $Q_b(n)^\dagger$  on the registers (C, R), accepting if (C, R) return to all  $|0\rangle$ 's state.

We are next to define the hiding (or concealing) and the binding properties of the scheme  $\{(Q_0(n), Q_1(n))\}_{n \geq 1}$ .

- **Statistically hiding.** We say that the scheme is statistically hiding if the quantum state ensembles  $\{\rho_0(n)\}_{n \geq 1}$  and  $\{\rho_1(n)\}_{n \geq 1}$  are quantum statistically indistinguishable.
- **Computationally  $\epsilon(n)$ -binding.** We say that the scheme is quantum computationally  $\epsilon(n)$ -binding if for any state  $|\psi\rangle$  in auxiliary register Z, and any efficiently realizable unitary transformation  $U$  performing on (R, Z),

$$\left\| (Q_1 |0\rangle \langle 0| Q_1^\dagger)^{CR} U^{RZ} ((Q_0 |0\rangle \langle 0|)^{CR} |\psi\rangle^Z) \right\| < \epsilon(n), \quad (2)$$

By the *reversibility* of quantum computation, the binding property can also be equivalently defined by swapping the roles of  $Q_0$  and  $Q_1$  in the above. Then the inequality (2) becomes

$$\left\| (Q_0 |0\rangle \langle 0| Q_0^\dagger)^{CR} U^{RZ} ((Q_1 |0\rangle \langle 0|)^{CR} |\psi\rangle^Z) \right\| < \epsilon(n). \quad (3)$$

We call  $\epsilon(n)$  the *binding error*. When  $\epsilon(n)$  is some negligible function, we usually drop it and just say that the scheme is computationally binding.

**Remark.** The (computational) binding property stated in the definition above is actually the *honest-binding*, which is equivalent to the sum-binding w.r.t. a generic quantum bit commitment scheme [Yan20].

In the sequel, to simplify the notation we often drop the security parameter  $n$  and just write  $(Q_0, Q_1)$  to denote a generic quantum computationally-binding bit commitment scheme.

We will use the scheme  $(Q_0, Q_1)$  to commit a binary string in a bitwise fashion. Namely, the quantum circuit to commit a string  $s = s_1 s_2 \cdots s_m \in \{0, 1\}^m$  is given by

$$Q_s \stackrel{def}{=} \bigotimes_{i=1}^m Q_{s_i}, \quad (4)$$

which performs on  $m$  copies of the quantum register pair (C, R).

## 2.2 Modeling an attack of the sender of quantum commitments

Consider a running of a larger two-party protocol in which a generic quantum bit commitment scheme is used and the sender of quantum commitments is *malicious*. The other party who will be referred to as the receiver is *honest*. The sender is supposed to commit to a string in  $\{0, 1\}^m$  in a bitwise fashion at some moment, and later try to open the commitments in a way as determined by the larger protocol. Then the behavior of the sender can be modeled by  $(U, |\psi\rangle)$  such that:

1. The sender prepares the system  $(\mathbb{C}^{\otimes m}, \mathbb{R}^{\otimes m}, \mathbb{D}, \mathbb{Z})$  in the quantum state  $|\psi\rangle$  at the *end* of the commit stage, and sends the commitment registers  $\mathbb{C}^{\otimes m}$  to the receiver.
2. In the reveal stage, the sender first performs the *unitary* transformation  $U$  on the system  $(\mathbb{R}^{\otimes m}, \mathbb{D}, \mathbb{Z})$ , and then sends registers  $(\mathbb{R}^{\otimes m}, \mathbb{D})$  to the receiver. The register  $\mathbb{D}$  is supposed to contain the classical information indicating *which* quantum bit commitments are to open as *what* value, and  $\mathbb{R}^{\otimes m}$  are decommitment registers.

We have two remarks about the modeling as above:

1. We note that there might be other operations performed by both the sender and the receiver *between* the end of the commit stage and the beginning of the reveal stage within the larger protocol. But in many cases, this can be simulated by absorbing these operations and auxiliary states introduced into the operation  $U$  and the state  $|\psi\rangle$ , respectively. Anyway, in this work we just restrict to consider the modeling as above for simplicity.
2. In the second item above, we assume without loss of generality that *all* decommitment registers  $\mathbb{R}^{\otimes m}$  are sent to the receiver in the reveal stage, though sometimes only a proper subset of commitments are required to open<sup>10</sup>. We can do this because the receiver is *honest*; sending all decommitment registers will not affect the security against the sender. More detail is referred to [FUYZ20].

## 2.3 Blum's zero-knowledge protocol for Hamiltonian Cycle

Basically, Blum's protocol [Blu86] proceeds as follows: on input a graph  $G$  (assuming it is represented by its adjacency matrix) with  $n$  vertices:

1. The prover first chooses a random permutation  $\Pi \in S_n$ , where  $S_n$  consists of all permutations over the set  $\{1, 2, \dots, n\}$ . Then it commits to the graph  $\pi(G)$ , sending all  $n^2$  (quantum) bit commitments to the verifier.
2. Upon receiving the prover's commitments, the verifier tosses a random coin to obtain the challenge bit  $b \in \{0, 1\}$  and sends it to the prover.
3. If the challenge  $b = 0$ , then the prover sends the permutation  $\pi$  together with the decommitment registers for *all* bit commitments to the verifier. If the challenge  $b = 1$ , then the prover sends the location of a Hamiltonian cycle  $H$  together with the decommitment registers for the commitments of all edges of the cycle  $H$  to the verifier.
4. If the challenge  $b = 0$ , then the verifier accepts if all bit commitments are opened as  $\pi(G)$  successfully. If the challenge  $b = 1$ , then the verifier accepts if the  $H$  is a possible location of a Hamiltonian cycle and all commitments to the edges of  $H$  are opened as 1 successfully.

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<sup>10</sup>For example, consider a running of Blum's zero-knowledge protocol for the language Hamiltonian Cycle in which the cheating prover responds to the challenge 1 of the verifier.

### 3 Generalized quantum computational binding

In our definition of quantum computational binding (inequalities (2) and (3) within Definition 4), we quantify over all efficiently realizable unitary transformations that do not touch the commitment. In this section, we show that we can generalize the quantum computational binding property by additionally quantifying over all efficiently realizable projectors. (Recall that we call a *projector*  $\Pi$  efficiently realizable if its induced binary measurement  $\{\Pi, \mathbb{1} - \Pi\}$  is efficiently realizable).

We remark that such a generalization is introduced mainly for a technical reason. (Refer to the subsequent section for its application.) Intuitively, such a generalization is needed because a larger quantum protocol within which quantum bit commitments are used may involve not only unitary transformations but also (projective) measurements. Its proof is by a standard technique, namely, any efficient adversary that is not unitary can be converted in a straightforward manner into an efficient adversary that is unitary by introducing ancilla qubits. The proof also makes an essential use of the *arbitrariness* of the efficiently realizable unitary transformation  $U$  and the auxiliary input state  $|\psi\rangle$  in the definition of quantum computational binding (Definition 4).

**Lemma 5** *Inherit all notations in Definition 4. Let the operator  $\Gamma = U_k \Pi_k \cdots U_1 \Pi_1$  be an arbitrary alternation of efficiently realizable unitary transformations and projectors, where  $k \geq 1$  is an integer, and for each  $i$  ( $1 \leq i \leq k$ ) both the unitary transformation  $U_i$  and the projector  $\Pi_i$  perform on the quantum registers  $(R, Z)$ . If the inequality (2) holds, then*

$$\begin{aligned} \left\| (Q_1 |0\rangle \langle 0| Q_1^*)^{CR} \Gamma^{RZ} ((Q_0 |0\rangle)^{CR} |\psi\rangle^Z) \right\| &< \epsilon(n), \\ \left\| (Q_0 |0\rangle \langle 0| Q_0^*)^{CR} \Gamma^{RZ} ((Q_1 |0\rangle)^{CR} |\psi\rangle^Z) \right\| &< \epsilon(n). \end{aligned}$$

PROOF: We just prove the first inequality; the second one can be proved symmetrically.

By the definition of quantum computationally binding (Definition 4), the first inequality trivially holds when the operator  $\Gamma$  is a unitary transformation. To handle projectors, the basic idea is *simulation*: namely, each binary projective measurement  $\{\Pi_i, \mathbb{1} - \Pi_i\}$  ( $1 \leq i \leq k$ ) performing on the registers  $(R, Z)$  can be simulated by a unitary transformation  $V_i$  performing on registers  $(R, Z, Y_i)$  in the standard way, where the register  $Y_i$  is a single qubit register initialized in the state  $|0\rangle$ . Note that if the binary measurement  $\{\Pi_i, \mathbb{1} - \Pi_i\}$  is efficiently realizable, then so is  $V_i$ . Put it formally,

$$V_i^{RZY_i} (|0\rangle^{Y_i} (Q_0 |0\rangle)^{CR} |\psi\rangle^Z) = |1\rangle^{Y_i} \otimes \Pi_i^{RZ} ((Q_0 |0\rangle)^{CR} |\psi\rangle^Z) + |0\rangle^{Y_i} \otimes (\mathbb{1} - \Pi_i)^{RZ} ((Q_0 |0\rangle)^{CR} |\psi\rangle^Z).$$

For each index  $i$  ( $1 \leq i \leq k$ ) and each bit  $b \in \{0, 1\}$ , we introduce the projector

$$\Pi_i^b \stackrel{\text{def}}{=} \begin{cases} \Pi_i, & \text{if } b = 1 \\ \mathbb{1} - \Pi_i, & \text{if } b = 0 \end{cases},$$

and the register  $Y = (Y_1, \dots, Y_k)$ . Then

$$(U_k^{RZ} V_k^{RZY_k} \cdots U_1^{RZ} V_1^{RZY_1}) (|0^k\rangle^Y (Q_0 |0\rangle)^{CR} |\psi\rangle^Z) = \sum_{s \in \{0,1\}^k} |s\rangle^Y \otimes U_k \Pi_k^{s_k} \cdots U_1 \Pi_1^{s_1} ((Q_0 |0\rangle) |\psi\rangle).$$

Hence,

$$\begin{aligned}
& \left\| (Q_1 |0\rangle \langle 0| Q_1^*)^{CR} (U_k^{RZ} V_k^{RZY_k} \dots U_1^{RZ} V_1^{RZY_1}) (|0^k\rangle^Y Q_0 |0\rangle^{CR} |\psi\rangle^Z) \right\|^2 \\
&= \left\| \sum_{s \in \{0,1\}^k} |s\rangle^Y \otimes (Q_1 |0\rangle \langle 0| Q_1^*) U_k \Pi_k^{s_k} \dots U_1 \Pi_1^{s_1} ((Q_0 |0\rangle) |\psi\rangle) \right\|^2 \\
&= \sum_{s \in \{0,1\}^k} \left\| |s\rangle \otimes (Q_1 |0\rangle \langle 0| Q_1^*) U_k \Pi_k^{s_k} \dots U_1 \Pi_1^{s_1} ((Q_0 |0\rangle) |\psi\rangle) \right\|^2 \\
&\geq \left\| |1^k\rangle \otimes (Q_1 |0\rangle \langle 0| Q_1^*) U_k \Pi_k^1 \dots U_1 \Pi_1^1 ((Q_0 |0\rangle) |\psi\rangle) \right\|^2 \\
&= \left\| (Q_1 |0\rangle \langle 0| Q_1^*) U_k \Pi_k \dots U_1 \Pi_1 ((Q_0 |0\rangle) |\psi\rangle) \right\|^2 \\
&= \left\| (Q_1 |0\rangle \langle 0| Q_1^*) \Gamma((Q_0 |0\rangle) |\psi\rangle) \right\|^2.
\end{aligned}$$

While the r.h.s. of the inequality above is exactly what we want to bound, the l.h.s. can be bounded by  $\epsilon^2$  due to the inequality (2) (which holds for any efficiently realizable unitary transformation and any auxiliary input state, in particular the unitary  $U_k V_k \dots U_1 V_1$  and the state  $|0^k\rangle |\psi\rangle$ , respectively). Then take the square root of both sides will finish the proof of the first inequality. ■

A straightforward corollary of the lemma above in the following will also be useful in our security analysis.

**Corollary 6** *Inherit all notations in Definition 4. Operator  $\Gamma$  is the same as introduced in Lemma 5. Quantum states  $|\psi_0\rangle, |\psi_1\rangle$  are two possible states of the register  $Z$ . Then*

$$\left| \left( \langle 0| Q_1^\dagger \rangle^{CR} \langle \psi_1 |^Z \right) \Gamma^{RZ} \left( (Q_0 |0\rangle)^{CR} |\psi_0\rangle^Z \right) \right| < \epsilon(n).$$

PROOF:

$$\left| \left( \langle 0| Q_1^\dagger \rangle^{CR} \langle \psi_1 |^Z \right) \Gamma^{RZ} \left( (Q_0 |0\rangle)^{CR} |\psi_0\rangle^Z \right) \right| \leq \left\| (Q_1 |0\rangle \langle 0| Q_1^*)^{CR} \Gamma^{RZ} \left( (Q_0 |0\rangle)^{CR} |\psi_0\rangle^Z \right) \right\| < \epsilon(n).$$

■

## 4 The predicate-binding property of quantum string commitments

In this section, we first introduce the notion of pattern-predicate and then the predicate-binding property of quantum string commitments. Next, we show that the parallel composition of a generic quantum computationally-binding bit commitment scheme gives rise to a quantum string commitment scheme that is predicate-binding w.r.t. a pair of inconsistent pattern-predicates of a special form. Last, we extend this predicate-binding property to a setting that is sufficient for our application, i.e. quantum zero-knowledge arguments for **NP**.

### 4.1 Pattern-predicate

Informally, the pattern-predicate defined in the below states that for a string to satisfy some predicate, it should exhibit a certain “pattern” somewhere. The intuition underlying our definition is that in typical applications of bit commitments, the receiver will check whether the value of the opened commitments will cause it to accept.



**Definition 7 (Pattern-predicate)** A *pattern-predicate*  $P$  on binary strings  $\{0,1\}^m$  ( $m \geq 1$ ) can be represented by a triplet of functions  $(\text{val}(\cdot), T(\cdot), s(\cdot))$ , where given a candidate *witness*  $w \in \{0,1\}^{\text{poly}(m)}$  as input:  $\text{val}(w) = 1$  if  $w$  is a valid witness, and 0 otherwise<sup>11</sup>;  $T(w)$  is a subset of  $\{1, 2, \dots, m\}$ ;  $s(w)$  is a string of length  $|T(w)|$ ; all three functions  $\text{val}(\cdot)$ ,  $T(\cdot)$ , and  $s(\cdot)$  can be computed in  $\text{poly}(m)$  time. A string  $str \in \{0,1\}^m$  *satisfies* the predicate  $P$  if there exists a (valid) witness  $w \in \{0,1\}^{\text{poly}(m)}$  satisfying  $\text{val}(w) = 1$  and  $str[T(w)] = s(w)$ , where  $str[T(w)]$  denotes the substring obtained from the string  $str$  by projecting it on coordinates in the subset  $T(w)$ .

**Remark.** Intuitively, a *valid* witness  $w$  for a string  $str$  guides us to find a pattern  $s(w)$  locating at positions specified by  $T(w)$  efficiently. This pattern will certify that the string  $str$  satisfies the pattern-predicate  $P$ . However, it might be *computationally hard* to find a valid witness for a given string  $str$ .

In this work, for simplicity we often drop the prefix “pattern” and just write “predicate” to refer to a pattern-predicate. For a predicate  $P$ , it induces a subset  $P$  (by abusing the notation) of strings in  $\{0,1\}^m$  such that a string  $s \in P$  if and only if it *satisfies* the predicate  $P$ ; we will identify a predicate as the subset induced by it. We say that two predicates  $P_1, P_2$  on the set  $\{0,1\}^m$  are *inconsistent* if  $P_1 \cap P_2 = \emptyset$ ; that is, no strings in  $\{0,1\}^m$  can satisfy both  $P_1$  and  $P_2$  simultaneously.

In a typical application of commitments within a larger protocol, at some stage of this protocol the party who plays the role of the possibly *cheating* sender of commitments will open commitments, and the party who plays the role of the *honest* receiver of commitments will do some verification. We note that it is this verification that naturally induces a pattern-predicate. See the following example.

**Example 1.** Consider a running of Blum’s zero-knowledge protocol for the **NP**-complete language Hamiltonian Cycle, in which the verifier is *honest* while the prover might be *cheating*, and the common input graph  $G$  has  $n$  vertices. Let  $m = n^2$ . Each graph with  $n$  vertices can be represented by an  $m$ -bit string. This running of Blum’s protocol induces two predicates on strings over  $\{0,1\}^m$ , corresponding to the verifier’s verifications w.r.t. two possible challenges, respectively. In more detail, when the verifier’s challenge is 0, it will check that all bit commitments are opened as a graph that is isomorphic to the input graph. This induces a predicate  $P_0$  which consists of all graphs that are isomorphic to the input graph. Formally, the predicate  $P_0$  can be represented by a triplet of functions  $(\text{val}(\cdot), T(\cdot), s(\cdot))$  such that: given a claimed permutation  $\pi$  over  $\{1, 2, \dots, n\}$ ,  $\text{val}(\pi) = 1$  if  $\pi$  indeed represents a valid permutation;  $T(\cdot) \equiv \{1, 2, \dots, m\}$ , and  $s(\pi) = \pi(G)$ . When the verifier’s challenge is 1, it will check that  $n$  (out of  $n^2$ ) bit commitments are opened as all 1’s; moreover, these  $n$  positions (of bit commitments to open) should correspond to a possible location of a Hamiltonian cycle. This induces a predicate  $P_1$  which consists of all  $n$ -vertices graphs containing a Hamiltonian cycle. Formally, the predicate  $P_1$  can be represented by a triplet of functions  $(\text{val}(\cdot), T(\cdot), s(\cdot))$  such that: given a claimed Hamiltonian cycle  $H$ ,  $\text{val}(H) = 1$  if  $H$  indeed represents a possible location of a Hamiltonian cycle;  $T(H)$  is set of coordinates corresponding to edges of  $H$ , and  $s(\cdot) \equiv 1^n$ . If the input graph is *not* Hamiltonian, then the two predicates  $P_0$  and  $P_1$  are obviously inconsistent.

Another example given below consider a simpler scenario, where a special form of pattern-predicates is introduced. In the sequel, we will study these special pattern-predicates first before more general ones.

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<sup>11</sup>Sometimes, it will be more convenient to identify the function  $\text{val}(\cdot)$  as an algorithm that decides the validity of a candidate witness.

**Example 2.** Consider the following scenario. The sender first commits to a string in a bitwise fashion. Later, *all* (bit) commitments are to open, and the receiver (of commitments) will check whether the whole revealed string satisfies an efficiently computable predicate  $P$  in the *common sense* (i.e. a predicate which can be evaluated on any input string in polynomial time, rather than pattern-predicate introduced in this work). Let  $A(\cdot)$  be an algorithm which runs in time  $\text{poly}(m)$  and can decide whether a string  $str \in \{0, 1\}^m$  satisfies  $P$ . We note that the predicate  $P$  can also be viewed as a pattern-predicate  $(A(\cdot), T(\cdot), s(\cdot))$  where  $T(\cdot) \equiv \{1, 2, \dots, m\}$  and  $s(\cdot)$  is the identity function; any string  $str \in P$  itself serves as its witness.

## 4.2 String predicate-binding

We first give an informal definition of the predicate-binding property of a quantum string commitment scheme, and then informally state we have achieved towards predicate-binding by composing a generic computationally-binding quantum bit commitment scheme in parallel. Last, we restate the definition of the predicate-binding w.r.t. the parallization of a generic computationally-binding quantum bit commitment scheme in a formal way for the purpose of proving predicate-binding in the sequel.

**Definition 8 (Predicate-binding, informal)** Let  $P_1, P_2$  be two *inconsistent* pattern-predicates. We say that a quantum string commitment scheme is *predicate-binding w.r.t.  $(P_1, P_2)$*  if any cheating sender, who can succeed in convincing the receiver that the committed value of the (claimed) quantum string commitment satisfies the predicate  $P_1$  with *certainty*, will fail to convince the receiver that the committed value satisfies the predicate  $P_2$  (except for a *negligible* probability). We say that a quantum string commitment scheme is *predicate-binding* if it is predicate-binding w.r.t. any pair of inconsistent predicates.

**Remark.** Classical commitments secure against classical attacks are trivially predicate-binding, simply because there is at most one string (i.e. the committed value) associated with each (claimed) commitment. However, this no longer holds w.r.t. either classical or quantum commitments secure against quantum attacks.

Restricting to consider the quantum string commitment scheme obtained by composing a generic computationally-binding quantum bit commitment scheme  $(Q_0, Q_1)$  in parallel, our goal is to show that it is predicate-binding w.r.t. inconsistent pattern-predicates pairs that are general enough for our applications (section 5). Informally, we can prove a theorem as below. We highlight (again) that we do not achieve the full predicate-binding, which is left as an interesting open problem.

**Theorem 2** *Suppose that the quantum bit commitment scheme  $(Q_0, Q_1)$  is computationally binding. Let  $P_1, P_2$  be two inconsistent predicates on the set  $\{0, 1\}^m$  such that for (at least) one of them, the verification of whether an  $m$ -bit string satisfies it needs to examine the bits at some fixed positions of the string (regardless of the witness provided). Then the parallel composition of the scheme  $(Q_0, Q_1)$  gives rise to a quantum string commitment scheme that is computationally predicate-binding w.r.t.  $(P_1, P_2)$ .*

For the purpose of proving Theorem 2, now let us restate Definition 8 w.r.t. the parallization of a generic computationally-binding quantum bit commitment scheme in a more formal way.

Suppose that a cheating sender who is modeled as in section 2.2 tries to convince the (honest) receiver that the committed value of a (claimed) quantum string commitment satisfies a predicate  $P = (\text{val}(\cdot), T(\cdot), s(\cdot))$ , i.e. the (claimed) commitment can be opened in such a way that if  $w$  is a

valid witness, then the bit commitments indexed by the subset  $T(w)$  are opened as the string  $s(w)$ . The predicate  $P$  naturally induces a *projector*  $P$  (also by abusing the notation) whose expression is given by

$$P = \sum_w (|w\rangle\langle w|)^D \otimes (Q_{s(w)}|0\rangle\langle 0|Q_{s(w)}^\dagger)^{C^{\otimes T(w)}R^{\otimes T(w)}}. \quad (5)$$

Its explanation follows. The summation is over all valid witnesses<sup>12</sup> for  $m$ -bit strings in  $P_1$ ; the quantum circuit  $Q_{s(w)}$  (whose meaning is referred to the equation (4)) performs on the copies of the quantum register pair  $(C, R)$  indexed by the subset  $T(w)$ ; in the reveal stage, the receiver will perform the *binary measurement*  $\{P, \mathbb{1} - P\}$  on its system to decide whether to accept or not. Hence, the sender's success probability of convincing the receiver to accept is given by  $\|PU|\psi\rangle\|^2$ , where recall that  $|\psi\rangle$  is the quantum state of the whole system at the end of the commit stage and  $U$  is the sender's operation in the reveal stage. We also note that the projector  $P$  is *efficiently realizable*, since all functions  $\text{val}(\cdot)$ ,  $T(\cdot)$  and  $s(\cdot)$  are efficiently computable.

Based on the expression (5), we can formalize the predicate-binding property of the parallelization of a generic quantum bit commitment scheme as follows.

**Definition 9 (Predicate-binding w.r.t. the parallel composition of QBC)** Let  $P_1, P_2$  be two *inconsistent* pattern-predicates. We say that the quantum string commitment scheme obtained by composing a generic quantum bit commitment scheme  $(Q_0, Q_1)$  in parallel is *predicate-binding w.r.t.  $(P_1, P_2)$*  if  $\|P_2UP_1|\psi\rangle\|^2$  is negligible, where  $|\psi\rangle$  is an arbitrary state of registers  $(C^{\otimes m}, R^{\otimes m}, D, Z)$ , and  $U$  could be any efficiently realizable unitary transformations that do not touch the quantum commitment (i.e. the commitment registers  $C^{\otimes m}$ ). We say that this quantum string commitment scheme is *predicate-binding* if it is predicate-binding w.r.t. any pair of inconsistent pattern-predicates.

In the remainder of this section, we will prove Theorem 2. We will first establish predicate-binding w.r.t. a special form of inconsistent pattern-predicate pair (as formalized in Lemma 10) and then extend it to the general case (as formalized in Theorem 3).

### 4.3 Towards predicate-binding: a special case

We first restrict to consider pattern-predicates arising in Example 2 in subsection 4.1, and try to establish predicate-binding w.r.t. such a pair of inconsistent predicates.

By instantiating the predicate  $P$  in the equation (5) with the predicate of the form introduced in Example 2, the expression of the projector  $P$  will become

$$P = \sum_{s \in P} (|s\rangle\langle s|)^D \otimes (Q_s|0\rangle\langle 0|Q_s^\dagger)^{C^{\otimes m}R^{\otimes m}}. \quad (6)$$

For any inconsistent predicate pair  $(P_1, P_2)$  whose corresponding projectors  $P_1$  and  $P_2$  are both of the form (6), we can prove the following main technical lemma of this work.

**Lemma 10** *Suppose that the scheme  $(Q_0, Q_1)$  is computationally  $\epsilon$ -binding for some arbitrary negligible function  $\epsilon(\cdot)$ . Both predicates  $P_1$  and  $P_2$  are of the form given by the expression (6). Then for any quantum state  $|\psi\rangle$  of registers  $(C^{\otimes m}, R^{\otimes m}, D, Z)$ , and any efficiently realizable unitary transformation  $U$  that does not touch the commitment registers  $C^{\otimes m}$ , we have  $\|P_2UP_1|\psi\rangle\|^2 \leq m^2\epsilon^2 + 2m\epsilon$ .*

<sup>12</sup>We point out that a string in  $P_1$  may have *multiple* witnesses.

PROOF: According to the expression (6), we can write

$$P_1 |\psi\rangle = \sum_{s \in P_1} \alpha_s |s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes m} R^{\otimes m}} \otimes |\phi_s\rangle^Z \quad (7)$$

$$= \sum_{s \in \{0,1\}^m} \alpha_s |s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes m} R^{\otimes m}} \otimes |\phi_s\rangle^Z, \quad (8)$$

where for  $s \notin P_1$ , we let  $\alpha_s = 0$  and  $|\phi_s\rangle$  be arbitrary; moreover, the complex coefficients  $\alpha_s$ 's satisfy  $\sum_{s \in \{0,1\}^m} |\alpha_s|^2 \leq 1$ . For convenience, we introduce the shorthand

$$|\psi_s\rangle \stackrel{\text{def}}{=} |s\rangle \otimes Q_s |0\rangle \otimes |\phi_s\rangle \quad (9)$$

for each  $s \in \{0,1\}^m$ . With these notations, our goal becomes to show

$$\left\| P_2 U \sum_{s \in \{0,1\}^m} \alpha_s |\psi_s\rangle \right\|^2 \leq m^2 \epsilon^2 + 2m\epsilon. \quad (10)$$

We will actually prove a strengthening of the inequality (10) by induction. Specifically, we will prove that for each  $k$  ( $0 \leq k \leq m$ ) and each string  $x \in \{0,1\}^{m-k}$ , it holds that

$$\left\| P_2 U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle \right\|^2 \leq (m^2 \epsilon^2 + 2k\epsilon) \sum_{s \in \{0,1\}^{k \circ x}} |\alpha_s|^2, \quad (11)$$

where  $\{0,1\}^k \circ x$  denotes the set of all  $m$ -bit strings with a suffix  $x$  of length  $m-k$ . If we view for each  $x \in \{0,1\}^{m-k}$ , where  $0 \leq k \leq m$ , it induces an internal node/leaf of a binary tree which corresponds to the summation  $P_2 U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle$ , then we will bound the (squared) norm of each internal node in a bottom-up way. Thus, the root of the tree will correspond to the case where  $k = m$  (then  $x$  becomes an empty string), i.e. l.h.s. of the inequality (10) without the squared norm. If we can prove the inequality (11), then plugging in  $k = m$  and the inequality  $\sum_{s \in \{0,1\}^m} |\alpha_s|^2 \leq 1$ , we will arrive at the inequality (10).

Now we are ready to prove the inequality (11) by induction on  $k$ , where  $0 \leq k \leq m$ .

Base. We show that the inequality (11) holds when  $k = 0$ . In this case,  $x$  is a string of length  $m$ . Since the coefficient  $\alpha_x = 0$  for  $x \notin P_1$ , in which case the inequality (11) holds trivially, we suffice to fix an arbitrary  $x \in P_1$  and show that  $\|P_2 U |\psi_x\rangle\| \leq m\epsilon$ . To this end, our technique is the *perturbation* that is similar to the quantum statistical binding setting [FUYZZ20]. Specifically, we will first show that the unit vector  $U |\psi_x\rangle$  is *negligibly close* to the (unnormalized) vector

$$|\tilde{\psi}_x\rangle \stackrel{\text{def}}{=} \bigotimes_{i=1}^m (\mathbb{1} - (Q_{\bar{x}_i} |0\rangle \langle 0| Q_{\bar{x}_i}^\dagger)) U |\psi_x\rangle, \quad (12)$$

where  $\bar{x}_i = 1 - x_i$ , and the projector  $Q_{\bar{x}_i} |0\rangle \langle 0| Q_{\bar{x}_i}^\dagger$  performs on the  $i$ -th copy of the register pair  $(C, R)$ . Second, we show that from the inconsistency of the predicate pair  $(P_1, P_2)$ , it follows that the vector  $|\tilde{\psi}_x\rangle$  is *orthogonal* to the subspace  $P_2$ . Combining these two facts, we know that  $\|P_2 U |\psi_x\rangle\|$  is negligible. Detail follows.

We first show that  $\|U |\psi_x\rangle - |\tilde{\psi}_x\rangle\| < m\epsilon$  via a simple hybrid argument. Specifically, we introduce hybrids for each  $0 \leq j \leq m$  such that  $H_j \stackrel{\text{def}}{=} \bigotimes_{i=1}^j (\mathbb{1} - Q_{\bar{x}_i} |0\rangle \langle 0| Q_{\bar{x}_i}^\dagger) U |\psi_x\rangle$ ; then  $U |\psi_x\rangle = H_0$  and

$|\tilde{\psi}_x\rangle = \mathbf{H}_m$ . We suffice to show that any two adjacent hybrids are negligibly close: if this is true, then applying the triangle inequality of the operator norm  $m$  times will yield the desired bound.

Indeed, for each  $1 \leq j \leq m$ ,

$$\begin{aligned}
& \|\mathbf{H}_j - \mathbf{H}_{j-1}\| \\
&= \left\| \bigotimes_{i=1}^j (\mathbf{1} - Q_{\bar{x}_i} |0\rangle\langle 0| Q_{\bar{x}_i}^\dagger) U |\psi_x\rangle - \bigotimes_{i=1}^{j-1} (\mathbf{1} - Q_{\bar{x}_i} |0\rangle\langle 0| Q_{\bar{x}_i}^\dagger) U |\psi_x\rangle \right\| \\
&\leq \left\| (\mathbf{1} - Q_{\bar{x}_j} |0\rangle\langle 0| Q_{\bar{x}_j}^\dagger) U |\psi_x\rangle - U |\psi_x\rangle \right\| \\
&= \left\| (Q_{\bar{x}_j} |0\rangle\langle 0| Q_{\bar{x}_j}^\dagger) U (|x\rangle Q_x |0\rangle |\phi_x\rangle) \right\| \\
&< \epsilon,
\end{aligned}$$

where the last “ $<$ ” follows from Lemma 5 by considering the  $j$ -th quantum bit commitment. In greater detail, to apply Lemma 5 we replace the  $|\psi\rangle$  and  $T$  in Lemma 5 with  $|x\rangle (\bigotimes_{i \neq j} Q_{x_i} |0\rangle) |\phi_x\rangle$  and  $U$  here, respectively.

We then show that the (unnormalized) vector  $|\tilde{\psi}_x\rangle$  is orthogonal to the subspace  $P_2$ , i.e.  $\|P_2|\tilde{\psi}_x\rangle\| = 0$ . This follows straightforwardly from the assumption that the predicate  $P_2$  is *inconsistent* with the predicate  $P_1$ . In greater detail, for each  $s \in P_2$ , we know that it is *different* from the string  $x \in P_1$ ; that is, there exists some index  $j$  ( $1 \leq j \leq m$ ) such that  $s_j = \bar{x}_j$ . Combining this with the equation (12), it follows that

$$\begin{aligned}
& \left\| (|s\rangle \langle s| \otimes Q_s |0\rangle\langle 0| Q_s^\dagger) |\tilde{\psi}_x\rangle \right\| \leq \left\| (Q_s |0\rangle\langle 0| Q_s^\dagger) |\tilde{\psi}_x\rangle \right\| \\
&\leq \left\| (Q_{\bar{x}_j} |0\rangle\langle 0| Q_{\bar{x}_j}^\dagger) \left( \bigotimes_{i=1}^m (\mathbf{1} - (Q_{\bar{x}_i} |0\rangle\langle 0| Q_{\bar{x}_i}^\dagger)) U |\psi_x\rangle \right) \right\| \\
&= 0.
\end{aligned}$$

Then summing over all  $s \in P_2$ , we obtain

$$\left\| \sum_{s \in P_2} (|s\rangle \langle s| \otimes Q_s |0\rangle\langle 0| Q_s^\dagger) |\tilde{\psi}_x\rangle \right\| = \|P_2|\tilde{\psi}_x\rangle\| = 0.$$

Combining  $\|U|\psi_x\rangle - |\tilde{\psi}_x\rangle\| < m\epsilon$  with  $\|P_2|\tilde{\psi}_x\rangle\| = 0$ , we arrive at  $\|P_2U|\psi_x\rangle\| \leq m\epsilon$ .

Induction. Now suppose that the inequality (11) holds for  $k-1$  and each binary string  $x$  of length  $m - (k-1)$ . We are to show that it also holds for  $k$  and an arbitrary binary string  $x$  of length of  $m - k$ .

For an arbitrary  $x \in \{0, 1\}^{m-k}$ , we first expand the l.h.s. of the inequality (11):

$$\begin{aligned}
& \left\| P_2U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle \right\|^2 = \left\| P_2U \sum_{s \in \{0,1\}^{k-1 \circ 0x}} \alpha_s |\psi_s\rangle + P_2U \sum_{s' \in \{0,1\}^{k-1 \circ 1x}} \alpha_{s'} |\psi_{s'}\rangle \right\|^2 \\
&\leq \left\| P_2U \sum_{s \in \{0,1\}^{k-1 \circ 0x}} \alpha_s |\psi_s\rangle \right\|^2 + \left\| P_2U \sum_{s' \in \{0,1\}^{k-1 \circ 1x}} \alpha_{s'} |\psi_{s'}\rangle \right\|^2 \\
&+ 2 \left| \sum_{s \in \{0,1\}^{k-1 \circ 0x}} \alpha_s \langle \psi_s | \cdot U^\dagger P_2 U \cdot \sum_{s' \in \{0,1\}^{k-1 \circ 1x}} \alpha_{s'} |\psi_{s'}\rangle \right|.
\end{aligned} \tag{13}$$

For convenience, we introduce shorthands

$$\alpha_{0x}^2 \stackrel{def}{=} \sum_{s \in \{0,1\}^{k-1} \circ 0x} |\alpha_s|^2, \quad \alpha_{1x}^2 \stackrel{def}{=} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} |\alpha_{s'}|^2, \quad \alpha_x^2 \stackrel{def}{=} \alpha_{0x}^2 + \alpha_{1x}^2.$$

Without loss of generality, we can assume that all  $\alpha_{0x}, \alpha_{1x}, \alpha_x \geq 0$ . With these notations, our goal (i.e. inequality (11)) becomes to show

$$\left\| P_2 U \sum_{s \in \{0,1\}^k \circ x} \alpha_s |\psi_s\rangle \right\|^2 \leq \alpha_x^2 (m^2 \epsilon^2 + 2k\epsilon),$$

and the induction hypothesis implies

$$\begin{aligned} \left\| P_2 U \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle \right\|^2 &\leq \alpha_{0x}^2 (m^2 \epsilon^2 + 2(k-1)\epsilon), \\ \left\| P_2 U \sum_{s \in \{0,1\}^{k-1} \circ 1x} \alpha_s |\psi_s\rangle \right\|^2 &\leq \alpha_{1x}^2 (m^2 \epsilon^2 + 2(k-1)\epsilon). \end{aligned}$$

The remainder of the analysis splits into two cases.

Case 1: either  $\alpha_{0x} = 0$  or  $\alpha_{1x} = 0$ . Without loss of generality, we can assume that  $\alpha_{1x} = 0$ . This implies that  $\alpha_{s'} = 0$  for each  $s' \in \{0,1\}^{k-1} \circ 1x$ . Thus,

$$\left\| P_2 U \sum_{s \in \{0,1\}^k \circ x} \alpha_s |\psi_s\rangle \right\|^2 = \left\| P_2 U \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle \right\|^2 \leq \alpha_{0x}^2 (m^2 \epsilon^2 + 2(k-1)\epsilon) \leq \alpha_x^2 (m^2 \epsilon^2 + 2k\epsilon),$$

where the first “ $\leq$ ” uses the induction hypothesis.

Case 2: both  $\alpha_{0x} > 0$  and  $\alpha_{1x} > 0$ . Following the inequality (13) and using the induction hypothesis, we have

$$\begin{aligned} \left\| P_2 U \sum_{s \in \{0,1\}^k \circ x} \alpha_s |\psi_s\rangle \right\|^2 &\leq \alpha_{0x}^2 (m^2 \epsilon^2 + (k-1)\epsilon) + \alpha_{1x}^2 (m^2 \epsilon^2 + 2(k-1)\epsilon) \\ &\quad + 2\alpha_{0x}\alpha_{1x} \cdot \underbrace{\left| \frac{1}{\alpha_{0x}} \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s \langle \psi_s | \cdot U^\dagger P_2 U \cdot \frac{1}{\alpha_{1x}} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle \right|}_{(*)}. \end{aligned}$$

We claim (refer to Claim 11 in the below) that the absolute value (\*) in the above can be bounded by  $2\epsilon$ . Then

$$\begin{aligned} \left\| P_2 U \sum_{s \in \{0,1\}^k \circ x} \alpha_s |\psi_s\rangle \right\|^2 &\leq (\alpha_{0x}^2 + \alpha_{1x}^2) (m^2 \epsilon^2 + 2(k-1)\epsilon) + 2\alpha_{0x}\alpha_{1x} \cdot 2\epsilon \\ &\leq (\alpha_{0x}^2 + \alpha_{1x}^2) (m^2 \epsilon^2 + 2(k-1)\epsilon) + (\alpha_{0x}^2 + \alpha_{1x}^2) \cdot 2\epsilon \\ &= \alpha_x^2 (m^2 \epsilon^2 + 2k\epsilon). \end{aligned}$$

The induction step is thus completed in both cases.

We finish the proof the inequality (11), and in turn the whole lemma. ■

We are left to prove the following claim.

**Claim 11** *The absolute value (\*) is less than  $2\epsilon$ .*

PROOF: Inherit all notations introduced within the statement and the proof of Lemma 10. Our idea is (again) using perturbation. Detail follows.

Plugging in the equation (9), *unit* vectors

$$1/\alpha_{0x} \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle \quad \text{and} \quad 1/\alpha_{1x} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle$$

can be written in the form

$$Q_0 |0\rangle \otimes |\xi_0\rangle \quad \text{and} \quad Q_1 |0\rangle \otimes |\xi_1\rangle,$$

respectively, where both  $Q_0 |0\rangle$  and  $Q_1 |0\rangle$  are the states of the  $k$ -th quantum register pair  $(C, R)$ , and

$$\begin{aligned} |\xi_0\rangle &= \frac{1}{\alpha_{0x}} \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |s\rangle \otimes Q_{s[\bar{k}]} |0\rangle \otimes |\phi_s\rangle, \\ |\xi_1\rangle &= \frac{1}{\alpha_{1x}} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |s'\rangle \otimes Q_{s'[\bar{k}]} |0\rangle \otimes |\phi_{s'}\rangle; \end{aligned}$$

here  $s[\bar{k}]$  and  $s'[\bar{k}]$  denote the substrings of  $s$  and  $s'$  with the  $k$ -th bit dropped, respectively.

We introduce more shorthands as follows:

$$\begin{aligned} |\eta_0\rangle &\stackrel{def}{=} U(Q_0 |0\rangle \otimes |\xi_0\rangle), & |\tilde{\eta}_0\rangle &\stackrel{def}{=} (\mathbb{1} - (Q_1 |0\rangle \langle 0| Q_1^\dagger))U(Q_0 |0\rangle \otimes |\xi_0\rangle), \\ |\eta_1\rangle &\stackrel{def}{=} U(Q_1 |0\rangle \otimes |\xi_1\rangle), & |\tilde{\eta}_1\rangle &\stackrel{def}{=} (\mathbb{1} - (Q_0 |0\rangle \langle 0| Q_0^\dagger))U(Q_1 |0\rangle \otimes |\xi_1\rangle), \end{aligned} \tag{14}$$

where both projectors  $\mathbb{1} - (Q_1 |0\rangle \langle 0| Q_1^\dagger)$  and  $\mathbb{1} - (Q_0 |0\rangle \langle 0| Q_0^\dagger)$  perform on the  $k$ -th quantum register pair  $(C, R)$ . With these notations, our goal becomes to show

$$|\langle \eta_0 | P_2 | \eta_1 \rangle| < 2\epsilon.$$

To this end, it suffices to show:

1.  $\| |\eta_0\rangle - |\tilde{\eta}_0\rangle \| < \epsilon;$
2.  $\| |\eta_1\rangle - |\tilde{\eta}_1\rangle \| < \epsilon;$
3.  $\langle \tilde{\eta}_0 | P_2 | \tilde{\eta}_1 \rangle = 0.$

This is because if all of the three items above hold, then a simple triangle inequality will finish the job.

Indeed, for the first item,

$$\| |\eta_0\rangle - |\tilde{\eta}_0\rangle \| = \left\| (Q_1 |0\rangle \langle 0| Q_1^\dagger)U(Q_0 |0\rangle \otimes |\xi_0\rangle) \right\| < \epsilon,$$

which follows immediately from the quantum computational binding property (Lemma 5).

Symmetrically, we can prove the second item.

For the third item, according to the equation (6),

$$\begin{aligned} P_2 &= \sum_{s \in P_2} (|s\rangle \langle s|)^D \otimes (Q_s |0\rangle \langle 0| Q_s^\dagger)^{C^{\otimes m} R^{\otimes m}} \\ &= \sum_{s \in P_2} (|s\rangle \langle s|)^D \otimes (Q_{s_k} |0\rangle \langle 0| Q_{s_k}^*) \otimes (Q_{s[\bar{k}]} |0\rangle \langle 0| Q_{s[\bar{k}]}^*), \end{aligned}$$

where the projector  $Q_{s_k} |0\rangle \langle 0| Q_{s_k}^*$  performs on the  $k$ -th quantum register pair (C, R), and the projector  $Q_{s[\bar{k}]} |0\rangle \langle 0| Q_{s[\bar{k}]}^*$  performs on the remaining  $m - 1$  copies of the quantum register pair (C, R). Whether  $s_k = 0$  or 1, following from the equations in (14),

$$\langle \tilde{\eta}_0 | (Q_{s_k} |0\rangle \langle 0| Q_{s_k}^\dagger) | \tilde{\eta}_1 \rangle = 0.$$

Thus,  $\langle \tilde{\eta}_0 | (|s\rangle \langle s| \otimes Q_{s_k} |0\rangle \langle 0| Q_{s_k}^\dagger) | \tilde{\eta}_1 \rangle = 0$ . Summing over all  $s \in P_2$  will yield  $\langle \tilde{\eta}_0 | P_2 | \tilde{\eta}_1 \rangle = 0$ .

This finishes the proof of the claim.  $\blacksquare$

**Towards a security reduction from (string) predicate-binding to quantum computational (bit) binding.** One can easily adapt our security analysis above to a construction of an attack of the quantum computational bit binding given access to an attack of the (string) predicate-binding. In greater detail, note that the only places we use the computational binding property of the scheme  $(Q_0, Q_1)$  lies in the base step and in bounding the absolute value (\*) within the induction step (Claim 11). Thus, if there were a successful attack  $(U, |\psi\rangle)$  of the predicate-binding property w.r.t. an inconsistent predicate pair  $(P_1, P_2)$ , then by contradiction there are two possibilities:

1. The inequality (11) fails for  $k = 0$ , some  $x \in \{0, 1\}^m$ , and the quantum state  $|\psi_x\rangle$ .
2. Claim 11 fails for some  $k$  where  $1 \leq k \leq m$ , some  $x \in \{0, 1\}^{m-k}$ , and quantum states  $1/\alpha_{0x} \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle$  and  $1/\alpha_{1x} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle$ .

Either of the two items above can be easily converted to an attack of the quantum computational binding property of the scheme  $(Q_0, Q_1)$ . But for this purpose, one needs to introduce both classical and quantum non-uniformities: classical non-uniformities include  $x$  in item 1 and  $k, x$  in item 2, while quantum non-uniformities include  $|\psi_x\rangle$  in item 1 and  $|\psi_s\rangle$  and  $|\psi_{s'}\rangle$  in item 2. This (hybrid) non-uniformity is similar to fixing “good” random coin tosses or a good index when invoking a hybrid argument in classical security reductions. In this sense, our security analysis above is not fully uniform.

#### 4.4 Extension

By slightly adapting its proof, we can extend Lemma 10 to hold for more general inconsistent predicate pairs so as to be useful in cryptographic applications. Specifically, we can prove Theorem 2. In the remainder of this subsection, we will first sketch why such an extension as described in Theorem 2 is possible, along the way setting up some notations. After this, we can state Theorem 2 in a more formal way (Theorem 3) before proving it.

**The generalization of the predicate  $P_2$ .** It is not hard to extend Lemma 10 to the case in which the predicate  $P_2$  is of the most *general* form as described by the equation (5) (while the restriction on the predicate  $P_1$  remains the same). This extension turns out to be already sufficient for our application (section 5). Now let us briefly mention how to adapt the proof of Lemma 10 to this case in the below.



Recall that the proof of Lemma 10 is by induction. For the base step, which essentially relies on the inconsistency of predicates  $P_1$  and  $P_2$ , almost the same proof as that of Lemma 10 goes through even if the predicate  $P_2$  is generalized. For the induction step, however, we will encounter new difficulty in bounding  $\langle \tilde{\eta}_0 | P_2 | \tilde{\eta}_1 \rangle$  within the proof of Claim 11 for the extension: now it may happen that some projectors of the form  $|w\rangle \langle w| \otimes (Q_{s(w)} |0\rangle \langle 0| Q_{s(w)}^\dagger)$  in the summation over all valid witnesses of the equation (5) do not touch the  $k$ -th register pair  $(C, R)$ . Thus, new technique is needed to handle such projectors for the purpose of bounding the absolute value (\*). Actually, this is where we really need to generalize the quantum computational binding property (refer to Definition 4) so as to cope with not only unitary transformations but also projectors (Lemma 5).

In greater detail, to bound the absolute value (\*) now we divide the summation over all valid witnesses in the equation (5) into two parts: the summations of those projectors that touch the  $k$ -th register pair  $(C, R)$ , and those do not. Correspondingly, we can bound the absolute value of each of these two parts separately, and then use the triangle inequality to get a bound of the absolute value (\*). In particular, the absolute value of the former part can be bounded by  $2\epsilon$  in a similar way as that of the proof of Claim 11, whereas the absolute value of the latter part can be bounded by  $\epsilon$ , thanks to the generalized computational binding property (Lemma 5). Combing them we obtain a  $3\epsilon$  bound of the absolute value (\*) in case of the generalized predicate  $P_2$ .

**The generalization of the predicate  $P_1$ .** Unfortunately, it seems unlikely that we can generalize the predicate  $P_1$  to the most general form (5) like the predicate  $P_2$  above by our technique. This is because the special form of the projector  $P_1$  (equation (6)) seem to play an important role in bounding the absolute value (\*) within the induction proof of Lemma 10. In more detail, it seems that we make an essential of the following *structure* of the superposition (8) (which is induced by the projector  $P_1$ ): for distinct  $s, s' \in \{0, 1\}^m$ , say  $s_i \neq s'_i$ , the projections of unit vectors  $|s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes m} R^{\otimes m}} \otimes |\phi_s\rangle^Z$  and  $|s'\rangle^D \otimes Q_{s'} |0\rangle^{C^{\otimes m} R^{\otimes m}} \otimes |\phi_{s'}\rangle^Z$  on the  $i$ -th quantum register pair  $(C, R)$  correspond to commitments to different bit value. It is this structure that makes our strategy of bounding the l.h.s. of the inequality (10) work.

In spite of the difficulty mentioned above, we still can generalize the predicate  $P_1$  to the case in which the associated function  $T(\cdot)$  is any *constant* function; that is, which bits are to examine for the verification of  $P_1$  are fixed. (In comparison, in the special form given by the equation (6), the function  $T(\cdot)$  is fixed to output the whole set and the function  $s(\cdot)$  is fixed to be the identity function.)

To have a glimpse of why such a generalization of  $P_1$  is possible, we first rewrite its expression in a proper form. Specifically, let  $T_1$  be the fixed subset that prescribes which bits are to examine for the verification of  $P_1$ . In this case whether a string  $s \in \{0, 1\}^m$  satisfies the predicate  $P_1$  actually only depends on its substring  $s[T_1]$ . Let  $l = |T_1|$ . The predicate  $P_1$  in turn induces a predicate  $P_1[T_1]$  on the set  $\{0, 1\}^l$  such that all  $l$  bits need to examine to check whether an  $l$ -bit string satisfies the predicate  $P_1[T_1]$ . Following the equation (5), the projector  $P_1$  can be written as

$$P_1 = \sum_w (|w\rangle \langle w|)^D \otimes (Q_{s(w)} |0\rangle \langle 0| Q_{s(w)}^\dagger)^{C^{\otimes T_1} R^{\otimes T_1}} \quad (15)$$

$$= \sum_{s \in P_1[T_1]} \sum_{w: s(w)=s} (|w\rangle \langle w|)^D \otimes (Q_s |0\rangle \langle 0| Q_s^\dagger)^{C^{\otimes T_1} R^{\otimes T_1}}. \quad (16)$$

Note that for distinct  $s, s' \in P_1[T_1]$ , the two projectors  $\sum_{w: s(w)=s} |w\rangle \langle w|$  and  $\sum_{w: s(w)=s'} |w\rangle \langle w|$

induce two subspaces that are *orthogonal*. Then similar to the equation (7), we can write

$$P_1 |\psi\rangle = \sum_{s \in P_1[T_1]} \alpha_s |\omega_s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes T_1} R^{\otimes T_1}} \otimes |\phi_s\rangle^{C^{\otimes(m-l)} R^{\otimes(m-l)} Z}, \quad (17)$$

where the unit vector  $|\omega_s\rangle$  is of the form  $\sum_{w:s(w)=s} \alpha_w |w\rangle^D$ , and vectors  $|\omega_s\rangle$  and  $|\omega_{s'}\rangle$  are orthogonal for distinct  $s, s' \in P_1[T_1]$ .

It is not hard to verify that if we replace this  $P_1 |\psi\rangle$  given by the equation (17) with the one given by the equation (7) in the proof of Lemma 10, then almost the same proof goes through.

**A formal statement of Theorem 2.** It turns out that predicates  $P_1$  and  $P_2$  in Lemma 10 can be generalized in the way as discussed above *simultaneously*. Thus, Theorem 2 can be stated in a more formal way as in the theorem below, whose proof is deferred to Appendix A.

**Theorem 3** *Suppose that the scheme  $(Q_0, Q_1)$  is computationally  $\epsilon$ -binding. Let  $P_1, P_2$  be two inconsistent predicates on the set  $\{0, 1\}^m$ , which induce two projectors of the form (16) and (5), respectively. Then for any quantum state  $|\psi\rangle$  of registers  $(C^{\otimes m}, R^{\otimes m}, D, Z)$ , and any efficiently realizable unitary transformation  $U$  that does not touch the commitment registers  $C^{\otimes m}$ , we have  $\|P_2 U P_1 |\psi\rangle\|^2 \leq m^2 \epsilon^2 + 3m\epsilon$ .*

## 5 Application: quantum zero-knowledge argument

In this section, we give an application of the quantum computationally predicate-binding string commitment scheme as shown in the proceeding section. Specifically, we show that Blum’s protocol for the NP-complete language Hamiltonian Cycle [Blu86] with a generic quantum computationally-binding bit commitment scheme plugged in gives rise to a quantum zero-knowledge *argument* system. While its quantum (perfect or statistical) zero-knowledge property can be obtained by a straightforward application of Watrous’s quantum rewinding technique [Wat09, Unr12, Unr16b, YWLQ15], its quantum computational soundness is established by Lemma 12 as stated below. Combing them we arrive at Theorem 1.

**Lemma 12** *Blum’s protocol for the language Hamiltonian Cycle with a generic quantum computationally-binding bit commitment scheme  $(Q_0, Q_1)$  plugged in is sound against any quantum provers who are polynomial-time bounded, with soundness error  $1/2 + \text{negl}(\cdot)$ .*

PROOF: This can be proved by instantiating Theorem 3 with proper predicates induced by Blum’s protocol. Detail follows.

Suppose that the binding error of the scheme  $(Q_0, Q_1)$  is  $\epsilon(\cdot)$ , which is a negligible function. We inherit notations as introduced in subsection 2.3. Following subsection 2.2, we can model a generic attack of the prover of Blum’s protocol in the following way. The combined (quantum) system of the (cheating) prover and the (honest) verifier is given by  $(P, D, C^{\otimes n^2}, R^{\otimes n^2})$ , where the  $n^2$  copies of the register pair  $(C, R)$  are used for (in total  $n^2$ ) quantum bit commitments; the register  $D$  will hold the classical information of the prover’s response (i.e. the permutation  $\pi$  when the challenge  $b = 0$  or the location of a Hamiltonian cycle  $H$  when  $b = 1$ ); the register  $P$  is the prover’s (private) workspace. Suppose that the whole system is initialized in the state  $|\psi\rangle$ . The prover sends the quantum register  $C^{\otimes n^2}$  to the verifier as its first message. Then depending on the challenge  $b$ , the prover will perform some polynomial-time realizable unitary transformation  $U_b$  on the registers  $(P, D, R^{\otimes n^2})$ . After receiving the prover’s response, the verifier will perform some

binary measurement, which also depends on the challenge  $b$  (as prescribed in the below), to decide to whether accept or not.

Formally, depending on the challenge  $b$ , the verifier's accepting conditions induce two pattern-predicates, which in turn induces two efficiently realizable projectors/binary measurements as follows:

1. The projector corresponding to  $b = 0$  is given by

$$P_0 = \sum_{\pi \in S_n} (|\pi\rangle \langle \pi|)^D \otimes (Q_{\pi(G)} |0\rangle \langle 0| Q_{\pi(G)}^\dagger)^{C^{\otimes n^2} R^{\otimes n^2}}.$$

2. The projector corresponding to  $b = 1$  is given by

$$P_1 = \sum_{H: n \text{ cycle}} (|H\rangle \langle H|)^D \otimes (Q_{1^n} |0\rangle \langle 0| Q_{1^n}^\dagger)^{C^{\otimes H} R^{\otimes H}},$$

where the projector  $Q_{1^n} |0\rangle \langle 0| Q_{1^n}^\dagger$  performs on the  $n$  copies of the register pair  $(C, R)$  that are determined by the location of the Hamiltonian cycle  $H$ .

We highlight that here we implicitly assume that the verifier just performs a big binary measurement (induced by either  $P_0$  or  $P_1$ ) to decide whether to accept or not; it in particular does not measure the register  $D$  to extract any classical information. It is easy to see that whether measuring the register  $D$  or not will not change the verifier's acceptance probability. But by doing this, we are then allowed to apply the quantum rewinding lemma (Lemma 2).

Now we are ready to argue the quantum computational soundness of Blum's protocol. Suppose for contradiction that there exists a efficiently realizable cheating prover given by  $(|\psi\rangle, U_0, U_1)$  as aforementioned who can break the quantum computational soundness. Namely,

$$\frac{1}{2} \sum_{b \in \{0,1\}} \|P_b U_b |\psi\rangle\|^2 > \frac{1}{2} + n^{-c},$$

where  $c$  is some constant. Then applying the quantum rewinding lemma (Lemma 2), it follows that

$$\left\| P_1 U_1 U_0^\dagger P_0 U_0 |\psi\rangle \right\| > n^{-c}. \quad (18)$$

On the other hand, we invoke Theorem 3 by doing the replacements as summarized in the following table:

Theorem 3	Blum's protocol
$m$	$n^2$
Registers $(C^{\otimes m}, R^{\otimes m})$	Registers $(C^{\otimes m}, R^{\otimes m})$
Register D	Register D
Register Z	Register P
Projector $P_1$	Projector $P_0$
Projector $P_2$	Projector $P_1$
Quantum state $ \psi\rangle$	Quantum state $U_0  \psi\rangle$
Unitary transformation $U$	Unitary transformation $U_1 U_0^\dagger$

In case that the input graph  $G$  is not Hamiltonian, the two predicates  $P_0$  and  $P_1$  are inconsistent. Applying Theorem 3 will yield an upper bound  $n^4 \epsilon^2 + 3n^2 \epsilon$  of the squared norm  $\|P_1 U_1 U_0^\dagger P_0 U_0 |\psi\rangle\|^2$ , which is negligible. But this contradicts with the inequality (18).

We finish the proof of the lemma. ■

## 6 Conclusion and open problems

In this work, we show that the parallel composition of a generic quantum computationally-binding bit commitment scheme gives rise to a quantum *string* commitment scheme that is computationally predicate-binding. This new notion of quantum computational string binding property is stronger than the trivial honest-binding property, and turns out to be useful in constructing quantum zero-knowledge arguments for **NP** languages. The main technical part of this work lies in establishing this quantum computational predicate-binding property.

There are many open problems following this work. In the below, we just mention some that interest us most:

1. Can we extend our technique to prove the predicate-binding w.r.t. inconsistent predicate pairs that are more general or with different *structure* requirements than that is stated in Theorem 2 (or formally Theorem 3)? Can we extend our technique to prove the predicate-binding w.r.t. multiple ( $\geq 3$ ) inconsistent predicates? If we can do this, then we may show that the GMW zero-knowledge protocol for the **NP**-complete language **Graph 3-Coloring** [GMW91] with a generic quantum computationally-binding bit commitment scheme plugged in gives rise to a quantum zero-knowledge argument. More generally, can we extend our technique to prove the  $Q$ -binding which is extended straightforwardly from the one for classical commitments secure against quantum attacks in [DFS04]?
2. Can we prove even stronger binding (than the predicate-binding) properties of the quantum string commitment scheme obtained by composing a generic quantum computationally-binding bit commitment scheme in parallel? Further, if this is possible, then can it yield any interesting applications? In particular, can we extend our technique to show  $f$ -binding introduced in [CDMS04] w.r.t. any interesting function  $f$  so that the corresponding quantum commitments could be useful in constructing quantum oblivious transfer like in [CDMS04] and beyond? We note that the answer to this question is affirmative in the setting where a generic quantum *statistically*-binding bit commitment scheme is composed in parallel [FUYZ20].
3. Can we show that plugging a generic quantum computationally-binding bit commitment scheme in a variant of Blum's protocol [Unr12, FUYZ20] gives rise to a quantum zero-knowledge argument-of-knowledge for the **NP**-complete language **Hamiltonian Cycle**?
4. How about plugging a generic computationally-binding quantum bit commitment scheme in [GK96] to obtain a quantum  $\epsilon$ -zero-knowledge proof in constant rounds like in [CCY20]? If this is true, then we can relax the complexity assumption required in [CCY20] to quantum-secure one-way functions by a quantum construction.

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## A A proof of Theorem 3

We just highlight how to adapt the proof of Lemma 10 to the setting of Theorem 3, in which both predicates  $P_1$  and  $P_2$  are generalized in the way as stated in section 4.4.

According to the equation (17), we can replace the equations (7), (8) in the proof of Lemma 10 with

$$\begin{aligned} P_1 |\psi\rangle &= \sum_{s \in P_1[T_1]} \alpha_s |\omega_s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes T_1} R^{\otimes T_1}} \otimes |\phi_s\rangle^{C^{\otimes(m-l)} R^{\otimes(m-l)} Z} \\ &= \sum_{s \in \{0,1\}^l} \alpha_s |\omega_s\rangle^D \otimes Q_s |0\rangle^{C^{\otimes T_1} R^{\otimes T_1}} \otimes |\phi_s\rangle^{C^{\otimes(m-l)} R^{\otimes(m-l)} Z}, \end{aligned}$$

where for  $s \notin P_1[T_1]$ , we let  $\alpha_s = 0$ , and the corresponding  $|\omega_s\rangle$  and  $|\phi_s\rangle$  be arbitrary<sup>13</sup>; moreover, the complex coefficients  $\alpha_s$ 's satisfy  $\sum_{s \in \{0,1\}^l} |\alpha_s|^2 \leq 1$ . We similarly introduce the shorthand

$$|\psi_s\rangle \stackrel{def}{=} |\omega_s\rangle \otimes Q_s |0\rangle \otimes |\phi_s\rangle,$$

and our goal becomes to show

$$\left\| P_2 U \sum_{s \in \{0,1\}^l} \alpha_s |\psi_s\rangle \right\|^2 \leq m^2 \epsilon^2 + 3m\epsilon.$$

We are to strengthen the inequality above and prove by induction that for each  $k$  ( $0 \leq k \leq l$ ) and each string  $x \in \{0,1\}^{l-k}$ , it holds that

$$\left\| P_2 U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle \right\|^2 \leq (l^2 \epsilon^2 + 3k\epsilon) \sum_{s \in \{0,1\}^{k \circ x}} |\alpha_s|^2, \quad (19)$$

Base. We first show that the inequality (19) holds when  $k = 0$ . In this case,  $x \in \{0,1\}^l$ . Since the coefficient  $\alpha_x = 0$  when  $x \notin P_1[T_1]$ , in which case the inequality (19) trivially hold, we suffice to fix an arbitrary  $x \in P_1[T_1]$  and show that  $\|P_2 U |\psi_x\rangle\| \leq l\epsilon$ .

It is not hard to see that the proof of the base step of Lemma 10 almost goes through here, except that now we apply the perturbation  $\bigotimes_{i=1}^l (\mathbb{1} - (Q_{\bar{x}_i} |0\rangle \langle 0| Q_{\bar{x}_i}^\dagger))$  to the subspace induced by the  $l$  copies of the quantum register pair  $(C, R)$  indexed by the subset  $T_1$ . In more detail, we introduce the (unnormalized) vector

$$|\tilde{\psi}_x\rangle \stackrel{def}{=} \bigotimes_{i=1}^l (\mathbb{1} - Q_{\bar{x}_i} |0\rangle \langle 0| Q_{\bar{x}_i}^\dagger) U |\psi_x\rangle,$$

which will play the same role of the equation (12) in the proof of the base step of Lemma 10. We remark that here we will make an essential use of the fact that predicates  $P_1$  and  $P_2$  are

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<sup>13</sup>We stress that our purpose of introducing  $\alpha_s, |\omega_s\rangle, |\phi_s\rangle$  for  $s \notin P_1[T_1]$  is mainly for a cleaner way of writing the proof; it will *not* affect the places in our proof where the (generalized) quantum computational binding property (Lemma 5) is applied.

*inconsistent*, so that the string  $s(w)$  w.r.t. each witness  $w$  within the expression of the projector  $P_2$  (refer to the equation (5)) must differ with the string  $x \in P_1[T_1]$  in at least one common coordinate  $i \in T_1 \cap T_2(w)$ .

Induction step. Now suppose that the inequality (19) holds for  $k - 1$  and each binary string  $x$  of the length  $l - (k - 1)$ . We are to show that it also holds for  $k$  and an arbitrary binary string  $x$  of the length of  $l - k$ .

Similar to the induction step of the proof of Lemma 10, now for each  $x \in \{0, 1\}^{l-k}$ , we similarly introduce shorthands

$$\alpha_{0x}^2 \stackrel{\text{def}}{=} \sum_{s \in \{0,1\}^{k-1} \circ 0x} |\alpha_s|^2, \quad \alpha_{1x}^2 \stackrel{\text{def}}{=} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} |\alpha_{s'}|^2, \quad \alpha_x^2 \stackrel{\text{def}}{=} \alpha_{0x}^2 + \alpha_{1x}^2,$$

where  $\alpha_{0x}, \alpha_{1x}, \alpha_x \geq 0$ . Then our goal becomes to show that

$$\left\| P_2 U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle \right\|^2 \leq \alpha_x^2 (l^2 \epsilon^2 + 3k\epsilon).$$

Indeed,

$$\begin{aligned} & \left\| P_2 U \sum_{s \in \{0,1\}^{k \circ x}} \alpha_s |\psi_s\rangle \right\|^2 = \left\| P_2 U \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle + P_2 U \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle \right\|^2 \\ & \leq \left\| P_2 U \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s |\psi_s\rangle \right\|^2 + \left\| P_2 U \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle \right\|^2 \\ & \quad + 2 \left| \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s \langle \psi_s | \cdot U^\dagger P_2 U \cdot \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle \right|. \end{aligned}$$

The remaining analysis also splits into two cases like that of the proof of Lemma 10, depending on whether at least one of  $\alpha_{0x}$  or  $\alpha_{1x}$  are zero. Now in the case that both  $\alpha_{0x} > 0$  and  $\alpha_{1x} > 0$ , we will encounter new difficulty in bounding the absolute value

$$\left| \frac{1}{\alpha_{0x}} \sum_{s \in \{0,1\}^{k-1} \circ 0x} \alpha_s \langle \psi_s | \cdot U^\dagger P_2 U \cdot \frac{1}{\alpha_{1x}} \sum_{s' \in \{0,1\}^{k-1} \circ 1x} \alpha_{s'} |\psi_{s'}\rangle \right|, \quad (20)$$

the counterpart of the absolute value (\*) within the proof of Lemma 10.

We will bound the expression (20) by  $3\epsilon$  in the following claim, which can be viewed as the counterpart of Claim 11. Once this is done, then we can complete the induction step similarly to that of the proof of Lemma 10 and establish Theorem 3.

We are left to prove the following claim.

**Claim 13** *The expression (20) is less than  $3\epsilon$ .*

**PROOF SKETCH:** We just highlight how to adapt the proof of Claim 11 to the setting here.

Compared with the proof of Claim 11, the new difficulty is: since now the projector  $P_2$  is of the most general form as given by the equation (5), it could happen that for some projector  $|w\rangle \langle w| \otimes Q_{s(w)} |0\rangle \langle 0| Q_{s(w)}^\dagger$  in the summation of  $P_2$ ,  $k \notin T(w)$ ; that is, this projector does not touch the  $k$ -th quantum register pair  $(C, R)$ . This will cause our argument for the equality  $\langle \tilde{\eta}_0 | P_2 | \tilde{\eta}_1 \rangle = 0$  within the proof of Claim 11 to fail.



To overcome this new difficulty, our idea is to split the projector  $P_2$  into two parts: the sum of projectors that touch the  $k$ -th quantum register pair  $(C, R)$ , i.e.  $k \in T(w)$ , which we denote by  $P_2^1$ , and the sum of those do not, which we denote by  $P_2^2$ . Then for the projector  $P_2^1$ , almost the same proof as that of Claim 11 will yield an upper bound  $2\epsilon$ , whereas for the projector  $P_2^2$ , we will use the generalized quantum computational binding property (Lemma 5) to obtain an upper bound  $\epsilon$ .

In more detail, after introducing similar notations  $|\xi_0\rangle, |\xi_1\rangle, |\eta_0\rangle, |\eta_1\rangle, |\tilde{\eta}_0\rangle, |\tilde{\eta}_1\rangle$  as in the proof of Claim 11, our goal is to show that  $|\langle \eta_0 | P_2 | \eta_1 \rangle| < 3\epsilon$ . Plugging in  $P_2 = P_2^1 + P_2^2$ , we will prove that

1.  $|\langle \eta_0 | P_2^1 | \eta_1 \rangle| < 2\epsilon$ , and
2.  $|\langle \eta_0 | P_2^2 | \eta_1 \rangle| < \epsilon$ .

For the item 1, by the property of the projector  $P_2^1$ , i.e. each projector in the summation of  $P_2^1$  touches the  $k$ -th quantum register pair  $(C, R)$ , almost the same proof as that of Claim 11 will yield the same upper bound  $2\epsilon$ .

For the item 2, the projector  $P_2^2$  does *not* touch the  $k$ -th quantum register pair  $(C, R)$ . We then apply Corollary 6, with the operator  $\Gamma$  replaced by  $UP_2^2U^\dagger$ , which will yield  $|\langle \eta_0 | P_2^2 | \eta_1 \rangle| < \epsilon$ . We additionally highlight that to apply Corollary 6, we need to show that the projector  $P_2^2$  is efficiently realizable given that the projector  $P_2$  is. This is indeed the case: conditioned on a quantum state collapsing to the subspace induced by the projector  $P_2$ , we can further compute the function  $T(w)$  given the witness  $w$  and check that  $k \notin T(w)$ .

Combining items 1 and 2 above, the absolute value (20) can be bound by  $3\epsilon$ . This finishes the proof of the claim. ■