Magic States Distillation Using Quantum LDPC Codes.

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1 Current Status.

1. Section 5 - Correct. In any CSS code, one can find a large subspace $\Lambda \subset C_X$ with a dimension that is linear in n and this subspace also satisfies the required relation for distillation. Specifically, for any $x \in \Lambda, y, z \in C_X$, it holds that $xy = 0$ and $xyz = 0$.

2. Sections 6 and 7 - Incorrect. Initially, I believed that assuming the code is LDPC, one could encode the state $C_Z^{\perp}$ in constant depth. However, this idea turned out to be incorrect both in calculation and in contrast to the fact that synthesizing the ground state of the Toric code requires $\Omega(\log n)$ depth.

2 Punctured Polyonomial Codes.

For $\Delta = 4^c$, $\Delta < q$, We have that

$$\sum_{x \in \mathbb{F}_\Delta} x^i =_\Delta \in \{0\ ,\ \Delta/2\}$$

$$\sum_{\substack{x \in \mathbb{F}_\Delta \\ x < \Delta}} x^{i+j} =_\Delta \sum_{x \in \mathbb{F}_\Delta} x^{i+j} =_\Delta \begin{cases} 0 & i+j \neq_q \Delta - 1 \\ \Delta - 1 & \text{else} \end{cases}$$

So the punctured d-dgree polynomial code is orthogonal for the punctured $n-1-d$ polyonimal code. So we can take $d = q/2 - 1$, and $\Delta = \alpha q$ to have $[\alpha q, q/2-1, q/2-(1-\alpha)q]$ code. For example we can take $\alpha = 7/8$ and have $[7/8q, q/2-1, 3/8q]$. The rate of the code is

$$\sim \frac{1}{2} / \frac{7}{8} = \frac{4}{7} > \frac{1}{2}$$

Claim 2.1. *For any $\Delta > 5$ there are good LDPC family C such that for any $x, y \in C$ it holds that $x \cdot y =_{(\Delta-1)} 0$.*

Proof. Consider the Tanner code defined by using the Δ-punctured polynomial code as C_0, where the rate of C_0 is strictly greater than $\frac{1}{2}$. Then we have for any $x, y \in C$:

$$x \cdot y =_{(\Delta-1)} \sum_{v \in V^+} x|_v \ \cdot \ y|_v =_{(\Delta-1)} 0$$

□

3 Candidate For Triorthogonal LDPC Code.

Consider the Tannner **Graph**, such that the graph G is bipartite, and every two checks overlap on the ith bucket, Δ-size, bits. So for any two checks, we have that

$$\begin{aligned}\sum_{x=i\cdot\Delta}^{(i+1)\Delta} x^j &=_\Delta \sum_{x'=(i-1)\cdot\Delta}^{i\Delta} (x'+\Delta)^j \\ &=_\Delta \sum_{x=(i-1)\cdot\Delta}^{i\Delta} x'^{,j} = \sum_{x\in\mathbb{F}_\Delta} x^j \\ \sum_{x\in\mathbb{F}_\Delta} (x+a\Delta)^i(x+b\Delta)^j &= \sum_{x\in\mathbb{F}_\Delta} x^{i+j}\end{aligned}$$

So it's left to show that if we take the bipartite graph to be an expander graph then we have a good code.

Let G be a bipartite graph $G=(L,R,E)$ that is a (n,m,γ,α) expander. This means that for any subset $S\subset V(G)$ with $|S|<\gamma n$, the size of the group of neighbors of S is at least $\Gamma(|S|)>\alpha|S|$. Consider the graph $G'=(\Delta\times L,R,E')$ defined as follows:

$$E'=\{\{(i,v),u\}:i\in[\Delta],\{u,v\}\in E\}$$

Thus for any $S\subset\Delta\times L$ if $|S|/\Delta<\gamma n$ we have that: $\Gamma'(S)<\Gamma(|S|/\Delta)$.

Therefore, if S is the set of vertices associated with the non-trivial symbols induced by the assignment of a codeword on the vertices, then if $|S|<\gamma n$, we have:

$$\frac{|S|}{\Gamma'(|S|)}\le\frac{|S|}{\Gamma(|S|/\Delta)}\le\frac{\Delta}{\alpha}$$

So there is a check that sees on his local view less than Δ/α non-trival bits $<d(C_0)$.

4 Hyprproduct Code of two Triorthogonal Codes.

Suppose that H is a parity check matrix scuh that $h_ih_j =_\Delta \in\{\Delta,\ ,\ \Delta/2\}$ for any two rows. IS that true that the same property holds for the following check matrix?

$$H'\leftarrow[H\otimes I|I\otimes H]$$

$$H'_iH'_j=(H\otimes I)_i(H\otimes I)_j+(I\otimes H)_i(I\otimes H)_j$$

Denote $i=(i_1,i_2)$ and $j=(j_1,j_2)$. So:

$$(H\otimes I)_i(H\otimes I)_j=\delta_{i_2,j_2}H_{i_1}H_{j_1}$$

and

$$(I\otimes H)_i(I\otimes H)_j=\delta_{i_1,j_1}H_{i_2}H_{j_2}$$

5 Good Codes With Large Λ.

Claim 5.1. *Let $v_1,v_2..v_k$ vectors in $\mathbb{F}_2^n$, then there are $u_1,u_2..u_{k'}$ for $k'>k/2$. Such span $\{u_1,u_2..u_{k'}\}\subset$ span $\{v_1,v_2..v_k\}$ and for any i,j it holds that $u_iu_j=0$.*

1 Let $J \leftarrow \emptyset$
2 **for** $i \in [k/2]$ **do**
3 $\quad J \leftarrow J \cup \{v_{2i-1}, v_{2i}\}$
4 $\quad$ **for** $S \subset J$ **do**
5 $\quad\quad$ Compute the vector m_S
6 $\quad\quad\quad$ define as $m_{S,j} = u_j \sum_{w \in S} w$
7 $\quad$ **end**
8 $\quad$ Pick S such $m_S = 0$ and set $u_i \leftarrow \sum_{w \in S} w$
9 $\quad$ Choose randomly $w \in S$ and set $J \leftarrow J/w$
10 **end**
: Find commuted vectors $u_1, u_2, ..u_{k'}$

1 Let $J \leftarrow \emptyset$
2 **for** $i \in [k/3]$ **do**
3 $\quad J \leftarrow J \cup \{v_{3i-2}, v_{3i-1}, v_{3i}\}$
4 $\quad$ **for** $S \subset J$ **do**
5 $\quad\quad$ Compute the vector m_S
6 $\quad\quad\quad$ define as $m_{S,j,j'} = u_{j'} u_j \sum_{w \in S} w$
7 $\quad$ **end**
8 $\quad$ Pick S such $m_S = 0$ and set $u_i \leftarrow \sum_{w \in S} w$
9 $\quad$ Choose randomly $w \in S$ and set $J \leftarrow J/w$
10 **end**
: Find commuted vectors $u_1, u_2, ..u_{k'}$

Proof. Consider Algorithm 1a, We are going to prove that at line number (8) the alg always finds a subset S that satisfies the equality. Assume not. On one hand, the number of possible values that m_S can have is $2^i - 1$. On the other hand, since J contains $i+1$ vectors on the ith iteration, it follows that the number of subsets is $2^{i+1} - 1 \geq 2^i$.

Therefore, there must be at least two different subsets S and S' such that $u_S = u_{S'}$. However, this means that

$$\begin{aligned} m_{S\Delta S',j} = u_j \sum_{w \in S\Delta S'} w &= u_j \left(\sum_{w \in S\Delta S'} w + 2 \sum_{w \in S \cap S'} w \right) \\ &= m_{S,j} + m_{S',j} = 0 \end{aligned}$$

Thus, $m_{S\Delta S'} = 0$. Additionally, it is clear that the rank does not decrease, as for u_i, there exists one v_j such that only u_i is supported by v_j. □

Claim 5.2. *Let $v_1, v_2..v_k$ vectors in $\mathbb{F}_2^n$ and m be an integer $m < k$, then there are $u_1, u_2..u_{k'}$ for $k' > k/2 - m$. Such span $\{u_1, u_2..u_{k'}\} \subset$ span $\{v_{m+1}, v_{m+2}..v_k\}$, for any i, j it holds that $u_i u_j = 0$ and for any $i \in]k'$, $j \leq m$ it holds that $u_i v_j = 0$.*

Proof. Modify the Algorithm 1a as follows, Initialize $u_1, .u_m$ to be $v_1, .., v_m$ and $J = \{v_{m+1}, ..v_{2m+2}\}$. Notice that in the ith iteration, for the counting argument to works in the proof of Claim 5.1, we have to ensure that:

$$\begin{aligned} |J| \geq& m + i + 1, \text{So } m + i + 1 \leq k - m - i \\ \Rightarrow& i \leq k/2 - m - \frac{1}{2} \end{aligned}$$

In the end, $u_{m+1}, u_{m+2}, .., u_{k'}$ will satisfy the equations. □

Claim 5.3. *Let $v_1, v_2..v_k$ vectors in $\mathbb{F}_2^n$, then there are $u_1, u_2..u_{k'}$ for $k' > k/4$. Such span $\{u_1, u_2..u_{k'}\} \subset$ span $\{v_1, v_2..v_k\}$. And for any i, j $\sum u_{i,k} u_{j,k} =_4 0$.*

Proof. Use the Algorithm 1a twice. However, in the second iteration, define $m_{S,j}$ to be the product of module 4. Note that $m_{S,j}$ must be either $4n$ or $4n+2$. Thus, we can follow the proof of Claim 5.1. □

Claim 5.4. *[COMMENT] Complete for the above the version, which handle triples. number of options is $(2^i)^2 = 2^{2i}$ and therefore we have the correctness if $|J| > 2i + 1$.*

Claim 5.5. *Consider the Left-Right (Δ,n)-Complex Γ. $\dim C_X/C_Z^{\perp} \cap C_Z/C_X^{\perp}$ is linear in n.*

Proof. The rates of both $C_X/C_Z^{\perp}$ and $C_Z^{\perp}/C_X^{\perp}$ are $(2\rho - 1)^2$, where ρ can be any number in the range $(0, 1)$ [LZ22]. Consider choosing ρ such that the rates of the quotient spaces are strictly greater than $\frac{1}{2} + \alpha$. This implies that the rate of their intersection is greater than 2α. □

Corollary 5.1. *Fix the rate of the small codes C_A and C_B to $\rho = \frac{1}{2} + \alpha$. There is a subspace $\Lambda \subset C_X/C_Z^{\perp}$ at rate $\frac{1}{4} \cdot 2\alpha$ such that for any $x \in \Lambda$ and $y, z \in C_Z^{\perp} \cup \Lambda$ it holds that:*

1. $xy =_4 0$
2. $xyz =_4 \sum_i x_i y_i z_i =_4 0$

Claim 5.6. *Consider C, Λ and C', Λ' defined in ??. Denote by $\bar{\Lambda}$ the subspace C/Λ. Then:*

$$d(C'/\bar{\Lambda}') \geq d(C/\bar{\Lambda})$$

Proof. The way we perform Guess elimination is critical. We want to make sure that we do not add an Λ row to a $\bar{\Lambda}$ row. **[COMMENT]** Continue, Easy. Just need to perform the row reduction when rows of Λ at bottom, and then rotate the matrix $\curvearrowright$

$$\begin{bmatrix} A & B \\ C & D \end{bmatrix} \curvearrowright \begin{bmatrix} D & C \\ B & A \end{bmatrix}$$

□

Claim 5.7 (Not Formal)**.** *It is easy to see that by using concatenation again, one can obtain the code* $\dim \Lambda' \leftarrow \frac{1}{2} \dim \Lambda'$. *For any $x \in$ gen Λ', $|x|_4 = 1$, and for any $x \in C'/\Lambda'$, we have $|x|_4 = 0$.*

Proof. **[COMMENT]** We will do it by iterating the generators of C after performing rows reduction to the generator matrix. Now we will concatenate the i coordinate to complete the weight of the ith row to satisfy the requirements. □

6 Compute $|C_Z^{\perp}\rangle$ In Constant Depth. [COMMENT] Wrong Section.

Let C_0 be a Δ-length error linear binary code, Γ a Δ-regular bipartite graph, and let C_Z be the Tanner code defined by C_0 and Γ. We are about to prove that the uniform superposition over $C_Z^{\perp}$ codewords can be computed with constant probability at a depth dependent only on Δ, in particular independent of the $C_Z^{\perp}$-length. For this, we are going to use Proposition 10 in [MN98], which states that both the encoder and the decoder of any stabilizer m-length code can be implemented by a circuit at depth $\Theta(\log m)$ with $\Theta(m^2)$ ancillae.

Claim 6.1. *Let G be a Δ-regular bipartite graph, and denote by $C_Z^{\perp}$ the dual-tanner code $\mathcal{T}(G, C_0^{\perp})^{\perp}$. Then there is a circuit that with constant probability computes the state $|C_Z^{\perp}\rangle$ at $\Theta(\log \Delta)$ depth, and $\Theta(\Delta^2)n$ ancillary qubits.*

Proof. Let E_v and D_v be the encoder and the decoder of C_0 over the local view of vertex v, By [MN98] we have that both have depth $\Theta(\log \Delta)$ and require Δ^2 ancillae. Since Γ is bipartite, we can decompose V into V^- and V^+ such that the local views of any two vertices in $V^{\pm}$ are disjoint. Therefore, for any two different vertices $v, u \in V^{\pm}$, the encoders E_v and E_u act on disjoint subsets of qubits, each corresponding to the local view of either v or u. Consider the following algorithm:

Initialize $2n$ qubits.
Call the left and right segments L and R.
Apply E_v in parallel on L for any $v \in V^+$.
Apply E_v in parallel on R for any $v \in V^-$.
XOR R into L by applying CNOT from the ith bit of R to the ith bit of L.
Apply D_v in parallel on R for any $v \in V^-$.
Apply H^k on L. And measure.
Accept if the result in C_Z

Algorithm 1: Compute $|C_Z^{\perp}\rangle$

For any $v \in V$, let $|z_v\rangle$ be the superposition of codewords in C_0 supported by the local view of v. Similarly, for any subset of vertices $W \subset V$, let $|z_W\rangle$ be the uniform superposition over the subspace spanned by the generators supported by the vertices in W. In other words:

$$|z_W\rangle = |\sum_{v \in W} z_v\rangle$$

Using the notation, applying the encoders E_v, E_u for any pair of vertices with disjoints local view become:

$$\begin{aligned} E_v \cup E_u |0\rangle^n &= E_v |0 + z_u\rangle = E_v |0_{/u\text{'s view}}\rangle \otimes |z_u\rangle \\ &= |z_v\rangle |z_u\rangle = |z_u + z_v\rangle = |z_{\{u,v\}}\rangle \end{aligned}$$

So applying all the encoders E_v at once over the positive vertices results in:

$$(\cup_{v \in V^+} E_v) |0\rangle^n = (\cup_{v \in V^+/v_0} E_v) |z_{v_0} + 0\rangle = |z_{V^+}\rangle$$

Thus the whole computation sum up into:

$$\begin{aligned} (\cup_{v \in V^+} E_v) \otimes (\cup_{v \in V^+} E_v) \qquad & |0\rangle^n \otimes |0\rangle^n \mapsto \\ \text{CNOT} \sum_{z \in A} \sum_{z' \in B} \qquad & |z_{V^+}\rangle |z_{V^-}\rangle \mapsto \\ I \otimes H^k \sum_{z \in A} \sum_{z' \in B} \qquad & |z + z'\rangle |z'\rangle \mapsto \\ \sum_{z \in A} \sum_{z' \in B} \qquad & |z + z'\rangle (-1)^{wz'} |w\rangle \mapsto \end{aligned}$$

So if $w \in C_Z$ then clearly $z'w = 0$. The probability for that to occur is

$$\mathbf{Pr}[w \in C_Z] = \frac{|C_Z|}{\mathbb{F}_2^n} = 2^{(\rho-1)n}$$

□

7 Distillate $|\Lambda + C_Z^{\perp}\rangle$ Into Magic.

Let $|f\rangle$ be a codeword in C_X, and let $\hat{X}_g$ be the indicator that equals 1 if f has support on generator g, and 0 otherwise. Observe that applying $T^{\otimes}$ on $|f\rangle$ yields the state:

$$\begin{aligned} T^{\otimes n} |f\rangle = T^{\otimes n} |\sum_g \hat{X}_g g\rangle &= \exp\Big(i\pi/4 \sum_g \hat{X}_g |g| - 2 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h |g \cdot h| \\ &+ 4 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| - 8 \cdot i\pi/4 \cdot \text{ integers} \Big) |f\rangle \\ &= \exp\Big(i\pi/4 \sum_g \hat{X}_g |g| - 2 \cdot \pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h |g \cdot h| + 4 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| \Big) |f\rangle \end{aligned}$$

So in our case:

$$\begin{aligned} T^{\otimes n} |f\rangle &= \\ &= \exp\Big(i\pi/4 \sum_{g \in \text{ gen } \Lambda} \hat{X}_g \\ &- 2 \cdot \pi/4 \sum_{g,h \in \text{ gen } C_Z^{\perp}} \hat{X}_g \hat{X}_h |g \cdot h| \\ &+ 4 \cdot i\pi/4 \sum_{g,h \in \text{ gen } C_Z^{\perp}} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| \Big) |f\rangle \end{aligned}$$

So eventually, we have a product of gates when non-Clifford gates are applied on only on generators of $C_Z^{\perp}$.

$$T^n |f\rangle = \prod_{g \in \text{ gen } \Lambda} T_g \prod_{g,h \in \text{ gen } C_Z^{\perp}} \{CS_{g,h}|CZ_{g,h}|I\} \prod_{g,h,l \in \text{ gen } C_Z^{\perp}} \{CCZ_{g,h,l}|I\} |f\rangle$$

Decompose $f = f_1 + f_2$, where f_1 is supported only on $C_X / C_Z^{\perp}$ and f_2 is supported only on $C_Z^{\perp}$. By using commuting relations, the above can be turned into.

$$\begin{aligned} T^n |f\rangle = &\prod_{g \in \text{ gen } \Lambda} T_g \ X_{f_1} \\ &\prod_{g,h \in \text{ gen } C_Z^{\perp}} \{CS_{g,h}|CZ_{g,h}|I\} \prod_{g,h,l \in \text{ gen } C_Z^{\perp}} \{CCZ_{g,h,l}|I\} |f_2\rangle \end{aligned}$$

Denote by M_1, M_2 the gates:

$$\begin{aligned} M_1 &= \prod_{g \in \text{ gen } \Lambda, h} \{CZ_{g,h}|I\} \\ M_2 &= \prod_{g,h \in \text{ gen } C_Z^{\perp}} \{CS_{g,h}|CZ_{g,h}|I\} \prod_{g,h,l \in \text{ gen } C_Z^{\perp}} \{CCZ_{g,h,l}|I\} \end{aligned}$$

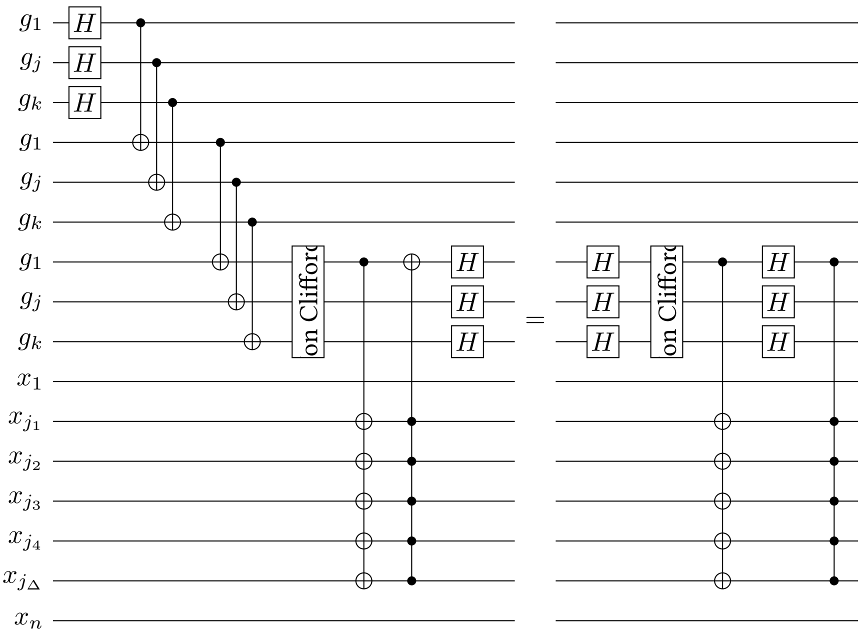
And then we get that

$$\begin{aligned} \prod_{g \in \text{ gen } \Lambda} T_g |f\rangle &= M_1^{\dagger} T^n M_2^{\dagger} |f\rangle \\ \prod_{g \in \text{ gen } \Lambda} T_g |f\rangle &= M_1^{\dagger} T^n \ \ E \ \ L[M_2^{\dagger}] \ \ |L[f]\rangle \end{aligned}$$

Claim 7.1. *Let $v \in V^{-}$, and let g_1 be the generator supported by v, which matches an assignment of a codeword in $C_A \otimes C_B$ on the local view of v. Denote by U_{v,g_1} the control-gate which, depending on the control bit $(v, 1)$, turns on g_1 over the edges associated with the local view of v in the graph G. Then, the depth of U_{v,g_1} depend only on Δ.*

Claim 7.2. *Let (v, g_1) and (u, g_2) be control wires for two different generators in the graph G. Then U_{v,g_1} and U_{u,g_2} **[COMMENT]** There must be a claim about the relationship between two different generators intersection, But I don't sure exactly why.*

Definition 7.1. *We say that a quantum circuit $\mathcal{C}$ is well error spreading if the light cone define by any T.*



Claim 7.3. *The state:*

$$\sum_{z \in C_Z^{\perp}} \exp\Big(-2 \cdot \pi/4 \sum_{g,h \in \text{ gen } C_Z^{\perp}} \hat{X}_g \hat{X}_h |g \cdot h| + 4 \cdot i\pi/4 \sum_{g,h \in \text{ gen } C_Z^{\perp}} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| \Big) |z\rangle$$

Can be computed such that any

Proof. Denote by U_v the gate which turn on all the generators supported on v. As any of them is just of a code word of $C_A \otimes C_B$, namely turning on generator require touching at most constant number of qubits combing □

Claim 7.4. *The state* $\left(M_2^{\dagger} \otimes I\right) |C_Z^{\perp} + \Lambda\rangle |0\rangle$ *can be computed, such that the light cone depth of any non-clifford gate is bounded by constant.*

Proof.

$$\begin{aligned}
(I \otimes H_X)\, CX_{n\to n}\, (E \otimes E) \quad I \otimes L[M_2^{\dagger}] \prod_{\substack{J \in \{\text{ gen } \Lambda, \\ \text{gen } C_Z^{\perp}\}}} \prod_{g \in J} (I + X_{L[g]}) \qquad & |0\rangle\,|0\rangle \\
= (I \otimes H_X)\, CX_{n\to n} \sum_{\substack{z \in C_Z^{\perp} \\ x \in \Lambda}} e^{\varphi(z)} \qquad & |x\rangle\,|z\rangle \\
= \sum_{\substack{z \in C_Z^{\perp} \\ x \in \Lambda}} e^{\varphi(z)} \qquad & |x+z\rangle\,|0\rangle \\
= \sum_{\substack{z \in C_Z^{\perp} \\ x \in \Lambda}} \left(M_2^{\dagger} \otimes I\right) \qquad & |x+z\rangle\,|0\rangle \\
= \left(M_2^{\dagger} \otimes I\right) \qquad & |C_Z^{\perp} + \Lambda\rangle\,|0\rangle
\end{aligned}$$

□

Denote by $p \in [0,1]$ the error rate of input magic states, and let $|A\rangle$ be an ancilla initialized to a one-qubit magic state. This $|A\rangle$ can be used to compute the T gate, with a probability of Z error occurring with a probability of p [BH12].

Claim 7.5. *There are constant numbers ζ_Δ, ξ_Δ, and a circuit $\mathcal{C}$ such that:*

1. *In the no-noise setting, The circuit compute the state*

$$\mathcal{C}\,|0\rangle^{\Theta(n)} \otimes |A\rangle^{\Theta(n)} \to \prod_{g \in \text{ gen } \Lambda} T_g \,|C_Z^\perp + \Lambda\rangle$$

2. *Otherwise, the circuit computes the state*

$$\mathcal{C}\,|0\rangle^{\Theta(n)} \otimes |A\rangle^{\Theta(n)} \to Z^e \prod_{g \in \text{ gen } \Lambda} T_g \,|C_Z^\perp + \Lambda\rangle$$

, where the probability that $e_i = 1$ is less than $\zeta_\Delta \cdot p$. Additionally, for any i, there are at most ξ_Δ indices j such that e_i and e_j are dependent.

Proof. Concatinate the $T^n \otimes I$ with the gate in Claim 7.4. ☐

Claim 7.6. *For any $\alpha \in (0,1)$ the probability that $|e| > (1+\alpha)p\zeta_\Delta$ is less than:*

$$\mathbf{Pr}\left[|e| > (1+\alpha)\mathbf{E}\left[|e|\right]\right] < \frac{1 \cdot \xi_\Delta n}{\alpha^2 \zeta_\Delta^2 p^2 n^2} = o\left(1/n\right)$$

Proof. By the Chebyshev inequality, notice that the number for which $\mathbf{E}\left[e_i e_j\right] - \mathbf{E}\left[e_i\right]\mathbf{E}\left[e_j\right] \neq 0$ is less than $\xi_\Delta n$. ☐

Definition 7.2. *We will said that a decoder $\mathcal{D}$ for the good qunatum LDPC code is an good-local decoder if*

1. *There is a treashold μn such that if the error size is less than $|e| < \mu n$ then $\mathcal{D}$ correct e in constant number of rounds. With probability $1 - o(1/n)$.*
2. *In any rounds $\mathcal{D}$ performs at most $O(n)$ work (depth $\times$ width).*
3. *The above is true in operation-noisy settings, where there is a probability of p for an error to occur after acting on a qubit. ($\star$)*

$\star$ *The motivation for this is that if the decoder does not act on the qubit, then it also does not apply a T gate on it. Therefore, in the distillation setting, there is zero chance for an error to occur.*

Claim 7.7. *Suppose there is a good local decoder $\mathcal{D}$ for the good qLDPC code. Then, there exists p_0 such that for any sufficiently large n, there is a distillation protocol that, given $\Theta(n)$ magic states at an error rate $p < p_0$, successfully distills $\Theta(n)$ perfect magic states with a probability of $1 - o(1/n)$. Furthermore, the protocol's space and time complexity (both quantum and classical) are $\Theta(n)$ and $\Theta(n^2)$, respectively.*

References

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