A Deterministic Polynomial-Time Protocol for Synchronizing from Deletions

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Abstract—In this paper, we consider a synchronization problem between nodes A and B that are connected through a twoway communication channel. Node A contains a binary file X of length n and node B contains a binary file Y that is generated by randomly deleting bits from X, by a small deletion rate β . The location of deleted bits is not known to either node A or node B. We offer a deterministic synchronization scheme between nodes A and B that needs a total of $O(n\beta \log \frac{1}{\beta})$ transmitted bits and reconstructs X at node B with probability of error that is exponentially low in the size of X. Orderwise, the rate of our scheme matches the optimal rate for this channel.

Keywords: Two-way communication, deletion channel, synchronization, edits, coding for synchronization.

I. INTRODUCTION

Consider two nodes A and B that respectively hold files X and Y, where file Y can be derived from file X by some deletions. For instance let

X = 0010110001010111, and Y = 00010001011.

Here Y is derived from X by 5 deletions, where deleted bits are denoted by D. We call Y a deleted version of X.

Suppose that the locations of deleted bits are *unknown* to both nodes. In this paper we are interested in the following question:

• What is the optimal transmission protocol for *synchronizing* the content of node *B* with the content of node *A*, i.e., how to reconstruct an estimate of file *X* at node *B*?

By way of optimality, we are mainly concerned with the number of transmitted bits between the two nodes and the complexity of implementing the protocol at nodes A and B. Also, as usual, we desire the reconstructed estimate of X at node B to have bit error probability that is exponentially small in the size of X.

Synchronization from deletions is a special case of a more general synchronization problem where file Y can be derived from X by a sequence of edits. An edit can refer to either deletion of a bit from file X or insertion of a new bit within X. File synchronization from random edits is the subject of many practical applications. Over the web, file updating is

an application where a user or a server needs to synchronize its outdated version of a file with a newer version. The new updates of a file can usually be modeled as random edits of its content. As another example, consider a search engine that constantly updates its database in order to reflect the latest changes to the content of websites. Here, as well, changes can be modeled by random edits to the content of websites. Another area of application is in distributed storage networks where several backup nodes store the same content and need to be regularly synchronized together. Mis–synchronization in storage devices can be due to mis-synchronized clock speeds of read and write heads of hard drives or crashes in random parts of the hard drive.

A. Previous Work

There has been a large body of research on synchronization from edits. In [1], Varshamov and Tenengolts offered a coding scheme for recovery from one asymmetric error. Soon thereafter, Levenshtein [2] showed that the scheme of Varshamov and Tenengolts can be used for synchronization from one deletion or one insertion. In [3], Orlitsky proved several fundamental bounds on the minimum number of transmitted bits under a restricted number of communication rounds for a prescribed edit distance. While the results of [3] are nonconstructive, several researchers have provided explicit code constructions. Let n denote the length of file X. For δ number of edits, Cormode et al. [4] offered an ϵ -error protocol with $c(\epsilon)\delta \log^3 n$ total transmitted bits¹, where $c(\epsilon)$ is a constant that depends on the error ϵ . For the same setting, Evfimievski [5] devised a protocol with the number of transmitted bits that is a polynomial in $\log n, \log \frac{1}{\epsilon}$, and δ . For an unknown, fixed number of edits δ , Orlitsky and Viswanathan [6] showed that the ϵ -error optimal protocol needs at most $\delta \log n + \log \frac{1}{\epsilon}$ transmitted bits. They also provided an explicit synchronization protocol that needs $2\delta \log n (\log n + \log \log n + \log \frac{1}{\epsilon} + \log \delta)$ transmitted bits. More recently, Venkataramanan et al. [7] offered a synchronization scheme that can correct $\delta = o(\frac{n}{\log n})$ edits with $(4c+1)\delta \log n$ transmitted bits from node A to node B and $10(\delta - 1)$ transmitted bits from node B to node A for any positive integer c. The error of reconstruction is at most $\frac{d \log n}{d k}$ where d is the number of deleted bits in X, out of δ total edits.

In practice, RSYNC [8] is a popular UNIX application for synchronizing between edited files. The RSYNC method can be in general very inefficient and the number of transmitted bits can be exponentially larger than the optimal number. There

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¹All logarithms in this paper are in base 2.

have been many improvements over the baseline approach. For example Suel *et al.* [9] proposed a protocol that in certain cases can save up to 50% of bandwidth over RSYNC. There are also more specialized synchronization tools, such as VSYNC [10], which synchronizes between video files.

B. Our Contribution

While most of the previous work has concentrated on synchronizing from a fixed number of edits between two files X and Y, in this paper we are interested in a more practical scenario, which is synchronizing from a *fixed rate* of edits between two files. We only study synchronization from deletions, and will discuss possible extensions to the more general case of deletions and insertions at the end of the paper. More specifically, we consider synchronization between node A and node B where node A has a binary string X that is generated by an i.i.d. Bernoulli process of parameter $\frac{1}{2}$. Node B has a binary string Y that is generated from X by randomly and independently deleting bits of X with probability β that is very small. We are interested in an optimal transmission protocol for synchronizing between nodes A and B when n, the length of X, is large.

We remark that, throughout the paper, by small β we implicitly mean that there exists $\beta_0 > 0$ such that our discussion is valid for all $\beta < \beta_0$. Furthermore, by large *n* we implicitly mean that for every $\beta < \beta_0$ there exists a positive integer n_β such that our discussion is valid for all $n > n_\beta$.

In order to evaluate a lower bound on the optimal number of transmitted bits between nodes A and B, suppose that node A has access to string Y. Then, the optimal number of transmitted bits to node B, needed for reconstructing X is H(X|Y), which is the conditional entropy of string X given string Y. Ma *et al.* [11] considered a more general set-up where the deletion pattern follows a stationary Markov chain. By applying the result of [11] to our model, for small values of β , the entropy H(X|Y) can be estimated as follows

$$H(X|Y) = n(\beta \log \frac{1}{\beta} + O(\beta)).$$
(1)

Therefore, any synchronization protocol needs at least $n(\beta \log \frac{1}{\beta} + O(\beta))$ transmitted bits. Paper [11] further uses tools from the well studied problem of *source coding with side information* [12], [13] to show that there exists a randomized synchronization protocol on a *one-way* channel that asymptotically needs H(X|Y) transmitted bits. However, [11] does not offer any explicit, deterministic construction for the synchronization protocol. We remark that the most efficient previous constructions (e.g., [7]) are for a *fixed* number of edits δ , and require $O(\delta \log n)$ transmitted bits between A and B. A naïve application of such results to our setup would require $O(n\beta \log n)$ transmitted bits between A and B for large n, which is clearly far from being optimal.

In this paper, we offer the first explicit and deterministic construction of a protocol for synchronizing from a small rate of deletions on a two-way, error-free channel. The protocol is optimal within a constant multiplicative factor and needs $O(n\beta \log \frac{1}{\beta})$ transmitted bits. Furthermore, we demonstrate that the error probability of synchronization at node *B* is

exponentially small in n. Finally, we show that our scheme needs a running time that is at most $O(n^4\beta^6)$.

The rest of the paper is organized as follows. In Section II, we present the problem setting and the main result along with a sketch of our synchronization scheme. In Section III, we present the mathematical details of our synchronization protocol and the proof of the main result in the paper. Section IV discusses practical implications of our protocol for low–complexity synchronization algorithms, and Section V includes concluding remarks and directions for possible extensions. Preliminary results from this work were reported in [14].

II. PROBLEM SETTING AND THE MAIN RESULT

A. Preliminaries

We represent a binary string Z of length ℓ by $Z = Z(1), Z(2), \dots, Z(\ell)$. For $1 \le i \le j \le \ell$, Z(i, j) denotes the substring $Z(i), Z(i+1), \dots, Z(j)$ of Z. If Z_1 is a string of length ℓ_1 and Z_2 is a string of length ℓ_2 , we denote by Z_1, Z_2 the string of length $\ell_1 + \ell_2$ obtained by concatenation of Z_1 and Z_2 . For a string Z, we let |Z| denote the length of Z.

Deletion channel is a channel that may delete any subset of the bits of the input string. Let X be the input to the deletion channel and Y be the output of the channel. We represent the set of deleted bits from X by a binary vector D of length |X|which is called the deletion pattern. If the deletion channel has deleted bit X(i) from X, then D(i) = 1 and otherwise D(i) = 0. For example, the output of a deletion channel with input X = 101 and deletion pattern D = 010, is Y = 11.

Corresponding to the deletion pattern D, we define a function f_D which maps the indices of bits in the input string, to their corresponding indices in the output string. If for index i, D(i) = 0, then $f_D(i) = i - \sum_{j < i} D(i)$, and if D(i) = 1, then $f_D(i) = f_D(i')$ where i' is the largest index, smaller than i, for which D(i') = 0. In the example above $f_D(1) = 1, f_D(2) = 1$, and $f_D(3) = 2$.

B. The Main Result

Suppose that node A contains a file that is represented by a binary string X of length n. Let node B contain file Y of length m that is the output of a deletion channel with input X and deletion pattern D. We assume that the deletion pattern is unknown to nodes A and B. Suppose that the source file X is generated by an i.i.d. Bernoulli source of parameter $\frac{1}{2}$ and that the deletion channel has deleted bits of X independently and with probability $\beta \ll 1$. We are interested in a synchronization protocol on a two-way, error-free channel between nodes A and B so that node B can recover string X from string Y with a small probability of error at the end of the communication session. Our main contribution in this paper is proving the following theorem.

Theorem 1. There exists a deterministic synchronization protocol between nodes A and B on a two-way, error-free channel, that on average transmits $O(n\beta \log \frac{1}{\beta})$ bits and

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generates an estimate $\hat{X} = \hat{X}(1), \dots, \hat{X}(n)$ of X at node B, such that $\Pr\left\{\hat{X}(i) \neq X(i)\right\} \leq 2^{-\Omega(n)}$ for every $1 \leq i \leq n$.

We prove the theorem by explicitly constructing a synchronization protocol. Next, we provide an overview of our synchronization protocol and prove its optimality.

C. Synchronization Protocol

Recall that node B has string Y which is a deleted version of string X. We next explain a synchronization protocol that enables node B to reconstruct an estimate of string X with a small probability of error. The synchronization protocol has three main steps, as illustrated in Figure 1. Each step is performed by a module at node B that has a two-way communication link to node A. The three modules work in series, such that the input to the first module is string Y and the output of the last module is the estimate \hat{X} of string X. Suppose that X is partitioned into substrings as follows

$$X = S_1, P_1, S_2, P_2, \cdots, S_{k-1}, P_{k-1}, S_k$$

where $|P_i| = L_P$ and $|S_i| = L_S$. Substrings P_1, \dots, P_{k-1} are called *pivot strings* and substrings S_1, \dots, S_k are called *segment strings*. We choose $L_S = \frac{1}{\beta}$ and $L_P = O(\log \frac{1}{\beta})$ and both node A and node B know the exact values of L_S and L_P . Note that the length of a pivot string is much smaller than the length of a segment string. We will determine the exact value of L_P later during our analysis.

1) The first step of the synchronization protocol is performed by the *matching module* at node B. In this step, node A sends pivot strings P_i , $1 \le i \le k-1$, in sequential order to node B. Upon receiving all the pivots, the matching module attempts to figure out the positions of pivots in Y by finding the exact copies of P_i 's within Y. The matching module is responsible for resolving ambiguities when there are multiple copies of a pivot in Y. The structure of the matching module and the graph-based algorithm for resolving the ambiguities are discussed in Section III. Due to possible deletions within P_i 's, the matching module is able to find the exact matches for only a subset of P_i 's. We will explain later the other possible cases when there are multiple matches for a pivot but an error is made by detecting a match that is not due to the original pivot. Suppose that the matching module finds matches for $P_{i_1}, \cdots, P_{i_{k'-1}}$ where $k' \leq k$. Based on the position of matched P_i 's, the matching module partitions Y into substrings as

$$Y = \bar{F}_1, P_{i_1}, \bar{F}_2, P_{i_2}, \cdots, \bar{F}_{k'-1}, P_{i_{k'-1}}, \bar{F}_{k'}$$

and sends this partitioned string to the next module, where \bar{F}_j denotes the substring between matched pivots $P_{i_{j-1}}$ and P_{i_j} in Y.

The next step is performed by the *deletion recovery* module at node B. After receiving the partitioned Y from the matching module, the deletion recovery module sends the indices {i₁, ..., i_{k'-1}} of the matched pivots in Y to node A. Upon receiving the indices, node A partitions X into substrings as follows:

$$X = F_1, P_{i_1}, F_2, P_{i_2}, \cdots, F_{k'-1}, P_{i_{k'-1}}, F_{k'}, \quad (2)$$

where F_j denotes the substring between pivots $P_{i_{j-1}}$ and P_{i_j} in X. Substring F_j can be written as follows:

$$F_j = S_{i_{j-1}+1}, P_{i_{j-1}+1}, \cdots, P_{i_j-1}, S_{i_j}.$$

Notice that if $P_{i_{j-1}}$ and P_{i_j} are matched correctly in Y, then \overline{F}_j can be derived from F_j by some sequence of deletions. In this step, nodes A and B use the synchronization protocol of Venkataramanan *et al.*, [7] with parameter c = 3 (c is a parameter that defines the tradeoff between complexity of the protocol and the error in the output of the decoder) to recover from deleted bits of \overline{F}_j and to form an estimate of F_j for each $1 \leq j \leq k$. Let us denote by \widetilde{F}_j the estimate of F_j at the output of the deletion recovery module. Notice that \widetilde{F}_j has the same length as F_j . At the end of this step, the deletion recovery module forwards the string

$$\tilde{X} = \tilde{F}_1, P_{i_1}, \tilde{F}_2, P_{i_2}, \cdots, \tilde{F}_{k'-1}, P_{i_{k'-1}}, \tilde{F}_{k'}$$
(3)

as an estimate of X to the last module.

3) At the last step, the LDPC decoder module at node B, recovers from the errors made by the first two steps. Due to a potential existence of multiple copies of each P_i within Y, the matching module (first step) may erroneously match P_i at a wrong place. Suppose P_{i_j} is a pivot that the matching module has matched at a wrong place. Then, \bar{F}_j and \bar{F}_{j+1} may not be realizable by deleting subsets of bits from F_j and F_{j+1} respectively. As a result, after the deletion recovery module (second step), \tilde{F}_j and \tilde{F}_{j+1} may be different from F_j and F_{j+1} , respectively. Furthermore, even if the matching module has matched pivots $P_{i_{j-1}}$ and P_{i_j} correctly in Y and \bar{F}_j is a deleted version of F_j , the protocol of Venkataramanan *et al.* [7], used in deletion recovery module, could introduce additional errors.

Suppose that the total error of the first two synchronization modules is bounded by ζ ,

$$\Pr\left\{\tilde{F}_j \neq F_j\right\} \leq \zeta.$$

We notice that the output of the deletion recovery module, \tilde{X} , is in synchronization with X, in the sense that $|\tilde{F}_j| = |F_j|$ for each $1 \le j \le k'$ and hence $\tilde{X}(i)$ is the estimate of X(i) for each index $1 \le i \le n$. Since the error rate over substrings \tilde{F}_j , $1 \le j \le k'$, is an upper bound for the bit error rate over X, we find that

$$\Pr\left\{\tilde{X}(i) \neq X(i)\right\} \le \zeta. \tag{4}$$

To recover from errors of \tilde{X} we use a powerful additiveerror correction code. Our choice is an LDPC decoder which receives parity check bits of a systematic LDPC code [15]. By applying a random permutation π at the input of the LDPC decoder and its inverse permutation π^{-1} at the output of the decoder, we can eliminate a potential non-uniformity of errors over different bits of \tilde{X} . Therefore, by using the error bound given in (4), the input sequence to the LDPC decoder can be modeled as an output of a Binary Symmetric Channel (BSC) with a Bernoulli i.i.d. input sequence of parameter $\frac{1}{2}$ and with



Figure 1. Illustration of the synchronization protocol.

the crossover probability of at most ζ . We assume that node A has access to the permutation π .

If node A sends a sufficient number of parity check bits to the LDPC decoder module, as shown in [16], the output of the decoder will be a string \hat{X} with

$$\Pr\left\{ \hat{X}(i) \neq X(i) \right\} \leq 2^{-\Omega(n)},$$

as previously stated in Theorem 1.

Next, we wish to estimate the total number of transmitted bits used by our synchronization protocol. We first establish a measure of the performance of the matching module of the decoder.

Theorem 2. Let $k' = (1 - L_P\beta + 2\beta + o(\beta))k$. For $L_P \ge 11 + 2\log \frac{1}{\beta}$, there exists a matching module that with probability $1 - 2^{-\Omega(n)}$, matches a subset $\{P_{i_1}, \dots, P_{i_{k'-1}}\}$ of pivots $\{P_1, \dots, P_{k-1}\}$ such that the probability of error in matching P_{i_i} is at most $\beta + o(\beta)$.

We devote Section III to proving this theorem. For the rest of our argument we set $L_P = 11 + 2\log \frac{1}{\beta}$, which is the minimum value of L_P required by Theorem 2.

Next, we use Theorem 2 to estimate the total number of transmitted bits needed by the synchronization protocol.

Lemma 1. On average, the total number of transmitted bits of the synchronization protocol is no more than $109n\beta \log \frac{1}{\beta}$.

Proof: First notice that $k = \frac{n+L_P}{L_S+L_P} = n\beta + 11\beta + 2\beta \log \frac{1}{\beta} = n\beta + o(1)$. The number of transmitted bits in the first step is

$$(k-1)L_P = 2n\beta \log \frac{1}{\beta} + o(n\beta \log \frac{1}{\beta}).$$

At the second step, node *B* needs no more than $k = n\beta$ bits to transmit the indices $\{i_1, \dots, i_{k'-1}\}$ to node *A*. Furthermore, the protocol of Venkataramanan *et al.* [7] for the recovery from deletions within each $F_j, 1 \leq j \leq k'$, with parameter c = 3 needs $13\delta_j \log |F_j| + 10(\delta_j - 1)$ transmitted bits, where

 $\delta_j := |F_j| - |F_j|$ is the number of deleted bits in F_j . Therefore, the average number of transmitted bits in the second step is no more than

$$n\beta + \mathbb{E}\left[\sum_{j=1}^{k'} \left(13\delta_j \log |F_j| + 10\delta_j\right)\right].$$

Notice that $\sum_{j=1}^{k'} \delta_j$ is the total number of deleted bits from X and is on average $n\beta$ (recall that we assumed that no deletions occurred in the matched pivots).

In Appendix I we show that $\mathbb{E} \left[\delta_j \log |F_j| \right] \le 16 + 8 \log \frac{1}{\beta}$. Therefore, the average number of transmitted bits in the deletion recovery module is upper bounded by

$$n\beta + k' \cdot 13(16 + 8\log\frac{1}{\beta}) + 10n\beta$$
$$\leq n\beta \cdot 13(16 + 8\log\frac{1}{\beta}) + 11n\beta$$
$$= 104n\beta \log\frac{1}{\beta} + o(n\beta \log\frac{1}{\beta}),$$

where we used the inequality $k' \leq k = n\beta + o(1)$.

For the last step, we would like to estimate the error ζ in \tilde{F}_j . By Theorem 2, the error probability in matching $P_{i_{j-1}}$ and P_{i_j} is at most $\beta + o(\beta)$ each. Since \bar{F}_j is the common neighbor of $P_{i_{j-1}}$ and P_{i_j} , with probability at most $2\beta + o(\beta)$, the string \bar{F}_j is not a deleted version of F_j . Furthermore, the error in the protocol of Venkataramanan *et al.* [7] for c = 3, is upper bounded by $\frac{\delta_j \log |F_j|}{|F_j|^3}$. Since $\mathbb{E}[\delta_j] = \beta L_S = 1$ and also $|F_j| = L_S = \frac{1}{\beta}$, the average probability of error by the protocol of Venkataramanan *et al.*, is upperbounded by $\beta^3 \log \frac{1}{\beta} = o(\beta)$. Counting the error from the matching module, we have $\Pr\left\{\tilde{F}_j \neq F_j\right\} \leq 2\beta + o(\beta)$, and therefore $\Pr\left\{\tilde{X}(i) \neq X(i)\right\} \leq 2\beta + o(\beta)$.

In order to recover from errors induced by a BSC with crossover probability of at most $2\beta + o(\beta)$, node A needs to send

$$nH(2\beta + o(\beta)) = 2n\beta \log \frac{1}{\beta} + o(n\beta \log \frac{1}{\beta}),$$

parity check bits to node *B*, where we use $H(\cdot)$ to refer to the binary entropy function defined as $H(t) = t \log \frac{1}{t} + (1 - t) \log \frac{1}{1-t}$ for 0 < t < 1.

The average number of transmitted bits in all three steps of the protocol is upper bounded by $108n\beta \log \frac{1}{\beta} + o(n\beta \log \frac{1}{\beta}) < 109n\beta \log \frac{1}{\beta}$. Therefore, the average number of transmitted bits by the algorithm is no more than $109n\beta \log \frac{1}{\beta}$.

In the next section we prove Theorem 2.

III. PROOF OF THEOREM 2

In this section, we propose a construction of a matching module such that for $L_P \geq 11 + 2\log \frac{1}{\beta}$, with probability $1 - 2^{-\Omega(n)}$, the module matches k' pivots, out of which at most βk pivots are matched erroneously. Since $\beta k = (\beta + o(\beta))k'$, our construction implies an error of at most $\beta + o(\beta)$ in matching the pivots. This claim is equivalent to the statement of Theorem 2.

We will frequently use the following concentration theorem in our argument:

Theorem 3 (Hoeffding [17]). Let p_0 be the probability that a biased coin shows heads. Then for every $\varepsilon > 0$, the probability that N tosses of the coin yield a number of heads between $(p_0 - \varepsilon)N$ and $(p_0 + \varepsilon)N$ is at least $1 - 2e^{-2\varepsilon^2N}$.

We will occasionally need a stronger version of the previous theorem:

Theorem 4 (Hoeffding [17]). Let z_1, \dots, z_N be i.i.d. random variables with expected value M that take values in an interval of length I. Then, for every $\varepsilon > 0$, the following holds

$$\Pr\left\{\left|\sum_{i=1}^{N} z_{i} - NM\right| \ge \varepsilon N\right\} \le 2\exp\left(-\frac{2\varepsilon^{2}N}{I^{2}}\right).$$

Recall that string X is partitioned into substrings as $X = S_1, P_1, \dots, S_{k-1}, P_{k-1}, S_k$, where $|S_i| = L_S$ and $|P_i| = L_P$. In our set-up, $L_S = \frac{1}{\beta}, L_P = O(\log \frac{1}{\beta})$, and $k = n\beta + o(1)$. Let us denote the index of the first bit of P_i in X by \check{p}_i and the index of the last bit of P_i in X by \hat{p}_i . Similarly, the first and last indices of S_i are denoted by \check{s}_i and \hat{s}_i . Therefore, $X(\check{p}_i, \hat{p}_i) = P_i$ and $X(\check{s}_i, \hat{s}_i) = S_i$.

The task of the matching module is to find "correct matches" of P_i 's within string Y. Next, we formalize the notion of correct and incorrect matches for a pivot P_i .

A. Correct and Incorrect Matches

Consider the substring $D(\check{p}_i, \hat{p}_i)$ which is the part of the deletion pattern D that acts on the pivot P_i . We consider the following cases:

- D(p̃_i, p̂_i) is the all zeros vector: There is no deletion within P_i. In this case we call the copy of P_i between indices f_D(p̃_i) and f_D(p̃_i) of Y the correct match of P_i. All other copies of P_i in Y are considered incorrect matches of P_i.
- D(p̃_i, p̂_i) has one nonzero element: There is one deletion within P_i. In this case, if there is a copy of P_i in Y that begins at f_D(p̃_i) or ends at f_D(p̂_i) then we call it a correct match of P_i and all other copies of P_i are called incorrect matches of P_i. If there is no such copy of P_i within Y, then all copies of P_i within Y are called incorrect matches. Notice that in this case there are possibly two correct matches for P_i. For instance, let P_i = 000 and let the immediate undeleted bits before and after P_i be zero. Then it is easy to verify that after one deletion within P_i, there is a copy of P_i starting at f_D(p̃_i) in Y and there is another copy of P_i ending at f_D(p̂_i) in Y.
- D(p̃_i, p̂_i) has more than one nonzero element: There is more than one deletion within P_i. In this case all copies of P_i within Y are considered incorrect matches.

While the definition of correct and incorrect matches is natural for the case of no deletion within P_i , we next explain the reason behind the definition for the case with deletions within P_i . Consider the illustration in Figure 2 where $P_i = 01101000$. Assume the penultimate bit is deleted from P_i . Suppose that



Figure 2. Illustration of a correct match of P_i with one deletion.

the bit right after P_i is 0. Notice that even with the deleted bit, a copy of P_i appears in Y, starting at $f_D(\check{p}_i)$. This copy of P_i is called a correct match. The reason is that the resulting string Y is the same as in the case where there is no deletion within P_i and instead the 0 after P_i is deleted in X. In other words, here we can "move" the deletion from P_i to the substring S_{i+1} without changing Y.

Although a similar scenario may happen when there are more than one deletions within P_i , i.e., we might be able to move the deleted bits from P_i to the neighboring segment strings without changing the resulting Y, since the probability of these cases is very small (the exact statement will follow), our analysis conservatively counts those matches as incorrect matches.

Next, we analyze the probability of occurrence of correct matches for P_i :

Lemma 2. With probability $1 - \beta L_P + o(\beta)$, P_i has no deletions and there is at least one correct match for P_i within Y.

Proof: With probability $(1 - \beta)^{L_P}$ no bit is deleted from P_i . For $L_P = O(\log \frac{1}{\beta})$ we have

$$(1-\beta)^{L_P} = 1 - L_P\beta + o(\beta).$$

Lemma 3. With probability $2\beta + o(\beta)$ there is one deletion within P_i and there is a correct match for P_i within Y.

Proof: Fix h as the place of the deleted bit out of L_P bits of P_i . Suppose $P_i(h) = b \in \{0, 1\}$. It is simple to observe that there is a copy of P_i starting at $f_D(\check{p}_i)$ in Y if and only if $P_i(h, L_P) = b, b, \cdots, b$ and furthermore, the first undeleted bit after P_i in X is also b. In other words, the hth bit of P_i should belong to the final "run" of zeros or ones of P_i and the first undeleted bit after P_i should also be of the same value. With probability $\beta(1 - \beta)^{L_P - 1}$, exactly the hth bit of P_i is deleted and with probability $2^{-(L_P - h + 1)}$ the bits after hth bit in P_i and the first undeleted bit after P_i have the same value as the *h*th bit of P_i . The overall probability of this case is $\beta(1-\beta)^{L_P-1}2^{-(L_P-h+1)}$. Similarly, there is a copy of P_i finishing at $f_D(\hat{p}_i)$ in Y if and only if all bits before the *h*th bit in P_i and the first undeleted bit before P_i are equal to the *h*th bit of P_i . This case happens with probability $\beta(1-\beta)^{L_P-1}2^{-h}$. The intersection of the two events happens when P_i is all-zeros or all-ones string and the immediate undeleted bits before and after P_i have the same value as the bits in P_i . This case happens with probability $\beta(1-\beta)^{L_P-1}2^{-(L_P+1)}$. By using the inclusion-exclusion principle and by varying *h* from 1 to L_P , we find the total probability of having one deletion within P_i , and as a result at least one correct match for P_i to be:

$$\beta(1-\beta)^{L_P-1} \sum_{h=1}^{L_P} \left(2^{-(L_P-h+1)} + 2^{-h} - 2^{-(L_P+1)} \right) = \beta(1-\beta)^{L_P-1} (2-2^{1-L_P} - L_P 2^{-(L_P+1)}) = 2\beta + o(\beta),$$

where in the last step we assumed $L_P = O(\log \frac{1}{\beta})$.

Lemma 4. With probability $o(\beta)$, P_i has more than one deletion.

Proof: Since the probability of no deletion within P_i is $(1 - \beta)^{L_P}$ and the probability of one deletion within P_i is $L_P\beta(1 - \beta)^{L_P-1}$, then the probability of more than one deletion within P_i is

$$1 - (1 - \beta)^{L_P} - L_P \beta (1 - \beta)^{L_P - 1} = o(\beta),$$

where we assumed $L_P = O(\log \frac{1}{\beta})$ in the final estimate. Let us define

$$R := 1 - L_P \beta + 2\beta. \tag{5}$$

From the preceding lemmas we conclude that:

Lemma 5. For a random string X and a random deletion pattern D, on average, the number of pivots with at least one correct match in Y is $(R + o(\beta))k$.

By applying Theorem 3 we conclude that:

Lemma 6. For a random string X and a random deletion pattern D, with probability $1 - 2^{-\Omega(n)}$, there are $(R + o(\beta))k$ pivots with at least one correct match in Y.

Proof: The probability that a pivot has a correct match in Y is $R+o(\beta)$ and it is independent of other pivots. Therefore, if in Theorem 3 we set p_0 to $R + o(\beta)$, N to k, and ε to $o(\beta)$, we conclude that the probability that for a random string X and a random string D there are between $(p_0 - o(\beta))k$ and $(p + o(\beta))k$ pivots with correct matches in Y, is at least $1 - 2e^{-2o(\beta)^2k} = 1 - 2^{-o(\beta)n} = 1 - 2^{-\Omega(n)}$. The fact that the set of integers between $(p_0 - o(\beta))k$ and $(p + o(\beta))k$ can be represented by the set of integers of the form $(R + o(\beta))k$, yields the result.

Lemma 7. For a random string X and a random deletion pattern D, with probability $1-2^{-\Omega(n)}$, there are $o(\beta)k$ pivots with two correct matches in Y.

Proof: As we showed in the proof of Lemma 3, the probability that a pivot has a deletion and two correct matches in Y is given by the following expression

$$L_P \beta (1-\beta)^{L_P-1} 2^{-(L_P+1)} = o(\beta),$$

where we assumed $L_P = O(\log \frac{1}{\beta})$. Therefore, the average number of pivots with two correct matches in Y is $o(\beta)k$. Now, if in Theorem 3 we set p_0 to $o(\beta)$, N to k, and ε to $o(\beta)$, we conclude that the probability that for a random string X and a random string D there are between $(p_0 - o(\beta))k$ and $(p + o(\beta))k$ pivots with two correct matches in Y, is at least $1 - 2e^{-2o(\beta)^2k} = 1 - 2^{-o(\beta)n} = 1 - 2^{-\Omega(n)}$. Since the set of integers between $(p_0 - o(\beta))k$ and $(p + o(\beta))k$ can be represented by the set of integers of the form $o(\beta)k$, the result follows.

B. The Matching Graph

The task of the matching module is to detect correct matches of P_i 's within Y. For this purpose we use a graph theoretic method. We define a graph G(V, E) with the vertex set as follows. Graph G has k+1 layers of vertices which are denoted by $\Lambda_0, \Lambda_1, \dots, \Lambda_k$. Each vertex in layer $\Lambda_i, 1 \le i \le k -$ 1, represents a match of pivot P_i in string Y. We refer to the vertices of Λ_i and matches of P_i in Y interchangeably. For vertex $v \in \Lambda_i$, let \check{v} and \hat{v} denote, respectively, the first and the last indices of the match of P_i corresponding to v in Y. We introduce two auxiliary vertices s and t where $\Lambda_0 =$ $\{s\}$ with $\hat{s} = 0$ and $\Lambda_k = \{t\}$ with $\check{t} = |Y| + 1$. Vertices s and t represent the beginning and the ending of string Y respectively.

We say a vertex in Λ_i is a *good* vertex if it corresponds to a correct match of P_i within Y. We call a vertex in Λ_i a *bad* vertex if it corresponds to an incorrect match of P_i . By definition of correct and incorrect matches, in each layer of graph G, there are possibly zero, one, or two good vertices. In order to detect the correct matches of P_i 's within Y, we need to find good vertices in graph G. For that, we define the edge set of G such that the good vertices are distinguished by their connectivity in the graph.

Let us define the distance between two vertices u and v in G as follows:

$$Dis(u, v) := \check{v} - \hat{u} - 1.$$

Notice that Dis(u, v) is nonnegative only when the first bit of v appears after the last bit of u. In that case, Dis(u, v) is the number of bits between u and v in Y.

For two pivots P_i and P_j with i < j in X, the number of bits between them in X is given by

$$(j-i-1)L_P + (j-i)L_S.$$

If both P_i and P_j have correct matches in Y, the number of bits between the correct match for P_i and the correct match for P_j is at most $(j - i - 1)L_P + (j - i)L_S$.

Furthermore, in most cases, for i < j, the first bit of the correct match for P_j appears after the last bit of the correct match for P_i . To see this, first notice that, since the first bit of P_j appears after the last bit of P_i in X, if there are no deletions within P_i and P_j , their order is preserved in Y.



Figure 3. Figure illustrates a graph G with 8 layers of vertices. The horizontal axis indicates different layers and the vertical axis indicates the position of each vertex in string Y that can take values from 1 to |Y|. The good and bad vertices are distinguished by black and white colors, respectively. The first layer has only one vertex s and the last layer has only one vertex t. As it is seen, all good vertices in the graph are connected together and they form an s - t path which is represented by the dashed edges in the graph.

Now consider the following example: let $P_i = 0000$ and $P_j = 0000$. Also assume that all bits between P_i and P_j are deleted except for a single 0 bit, and assume that exactly one bit is deleted from P_i and exactly one bit is deleted from P_j . In this case, the compound substring of Y corresponding to P_i and P_j and the bits in between them in X is 0000000, where the first four bits constitute the correct match for P_j . As we can observe, the first bit for the correct match of P_j is the last bit for the correct match for P_i and the correct match for P_i is the last bit for the correct match for P_j . It is easy to verify that in general for j > i, the least value of the distance between the correct match of P_j is -1.

Based on the two preceding observations, we connect a vertex $u \in \Lambda_i$ to a vertex $v \in \Lambda_j$ if and only if

$$-1 \le \text{Dis}(u, v) \le (j - i - 1)L_P + (j - i)L_S.$$
(6)

Therefore, all pairs of good vertices from different layers are connected together. By definition, s and t, which indicate the beginning and the ending of string X, respectively, are treated as "auxiliary" good vertices. Therefore, good vertices across different layers form an s - t path in graph G. However, there are potentially many other pairs of vertices that satisfy the condition of (6) and are connected together. Figure 3 illustrates an instance of graph G with 8 layers and the connections between vertices.

The following theorem shows that, with very high probability, bad vertices do not contribute to an s - t path. That is, any s - t path of the appropriate length in graph G is formed mostly of good vertices. Recall the definition of R from (5). We then have the following result. **Theorem 5.** Let X be a random input string to a deletion channel and D be a random deletion pattern. Let Y be the string obtained from X and D. Let G denote the matching graph corresponding to Y. Then, for $L_P \ge 11 + 2\log \frac{1}{\beta}$, with probability at least $1 - 2^{-\Omega(n)}$, all paths from s to t with $Rk + o(\beta)k$ vertices, have at least $Rk - \beta k + o(\beta)k$ good vertices.

Theorem 5 is not only an existence statement, but also has an algorithmic implication. The implication is that if we pick any path from s to t with $Rk + o(\beta)k$ vertices, the path has many good vertices. Since finding an s - t path of an appropriate length in G is a computationally tractable task (we will discuss the computational complexity in the next section), finding a large fraction of good vertices is also a tractable task.

Overview: Before presenting the detailed proof of the theorem we first sketch the overall idea of the proof. To prove the theorem, we show that for a random string X and a random deletion pattern D, the probability of the existence of an s-tpath Q in G with $Rk + o(\beta)k$ vertices, such that the number of good vertices on Q is less than $Rk - \beta k + o(\beta)k$ is upper bounded by $2^{-\Omega(n)}$. Equivalently, we show that the probability of the existence of an s-t path Q with $Rk + o(\beta)k$ vertices such that number of bad vertices on Q is more than $\beta k + o(\beta)k$ is upper bounded by $2^{-\Omega(n)}$. To find an upper bound on the latter probability we use the union bound: for every α with $\beta \leq \alpha \leq R + o(\beta)$, we find an upper bound on the probability that there exists an s - t path Q with $Rk + o(\beta)k$ vertices such that the number of bad vertices on Q is αk . Then, by integrating the upper bound over all values of α and showing that it is less than $2^{-\Omega(n)}$, we conclude the result.

For a fixed value of α we evaluate an upper bound on the probability of the existence of an s - t path Q with $Rk + o(\beta)k$ vertices and αk bad vertices in the following way. Let us denote all good vertices of G by U, where $|U| = Rk + o(\beta)k$ with probability at least $1 - 2^{-\Omega(n)}$. We fix the realizations of all \check{u} for which $u \in U$. In other words, we fix the positions of good vertices of graph G. For Q to have exactly αk bad vertices and $Rk - \alpha k + o(\beta)k$ good vertices, we first choose $Rk - \alpha k + o(\beta)k$ good vertices of Q from the set U (and account for the cases with possibly two correct matches). Graph G has k + 1 layers and $Rk - \alpha k + o(\beta)k$ have been chosen to include the good vertices of Q. The remaining αk vertices of Q are chosen from the remaining $k + 1 - (Rk - \alpha k + o(\beta)k)$ layers. Since the vertices in set U have fixed positions in Y, all good vertices of Q have fixed positions in Y. However, we have only fixed the layers which include the bad vertices of Q and not the positions of bad vertices in Y. Next, to find an upper bound on the number of possible positions of the bad vertices of Q, we use a combinatorial argument based on the constraints imposed by the connectivity of consecutive vertices on Q via the edges of graph G. We notice that the positions of all vertices on Q are uniquely determined based on the distances between consecutive vertices on Q. Since the good vertices on Q have fixed positions, the distance between two consecutive vertices on Q, where both of them are good vertices, is fixed. However, the distance of two consecutive vertices on Q, where one of the vertices is a bad vertex, is a variable. We need to find all solutions to these variables such that constraints defined by (6) are satisfied. We show how to consolidate all resultant edge constraints over all edges of Q into a single linear constraint, and then by counting the number of solutions to that constraint we find an upper bound on the number of possible positions of the bad vertices of Q.

Finally, we notice that the probability that a substring P_i has an incorrect match in Y at some specific position is 2^{-L_P} . Therefore, if we are given the positions of all bad vertices on Q, the probability that there are incorrect matches of the corresponding pivots at those positions is $2^{-\alpha kL_P}$ each. By multiplying $2^{-\alpha kL_P}$ by the upper bound on the possible number of positions for bad vertices of Q, we find an upper bound on the probability of the existence of Q with $Rk+o(\beta)k$ vertices and αk bad vertices. Next we present the details of our argument.

Proof: We begin by finding an upper bound on the probability of the existence of a path Q from s to t with $Rk + o(\beta)k$ vertices out of which αk are bad vertices, for some $\beta \leq \alpha \leq 1$. There are k + 1 layers in graph G and by Lemma 6 with probability $1 - 2^{-\Omega(n)}$ there are $Rk + o(\beta)k$ layers with good vertices in graph G. Let us fix the realization of the deletion pattern D, the realization of the pivots P_i in X with exactly one deletion, and the realization of the immediate undeleted bits before and after pivots P_i in X with exactly one deletion. In this way, good vertices of graph G are fixed. We consider two cases:

Case 1: $\beta \leq \alpha < \frac{1}{2}$

For $\beta \leq \alpha < \frac{1}{2}$, first we fix the layers which have a vertex on the path Q of length $Rk + o(\beta)k$. Since by assumption, there are $Rk - \alpha k + o(\beta)$ good vertices on the path Q, the selection of good and bad vertices on the path can be done in at most the following number of ways

$$\binom{Rk+o(\beta)k}{Rk-\alpha k+o(\beta)k} \cdot 2^{o(\beta)k} \cdot \binom{(1-R)k+\alpha k+o(\beta)k}{\alpha k} < 2^{k((R+o(\beta))H(\frac{\alpha}{R+o(\beta)})+o(\beta)+(1-R+\alpha+o(\beta))H(\frac{\alpha}{1-R+\alpha})))} = 2^{k(RH(\frac{\alpha}{R})+(1-R+\alpha)H(\frac{\alpha}{1-R+\alpha})+o(\beta))}.$$

$$(7)$$

In the multiplication above, the first term stands for the number of ways we can choose the layers with good vertices on the path Q. By Lemma 7, with probability $1 - 2^{-\Omega(n)}$, there are $o(\beta)k$ pivots with two correct matches in Y. Thus, with probability $1 - 2^{-\Omega(n)}$, there are at most $o(\beta)k$ layers with two good vertices among the layers with good vertices on Q. Therefore, the second term in the multiplication above, is an upper bound on the number of combinations we can pick one good vertex from each of those layers. The last term stands for the number of ways we can choose the layers with bad vertices from the remaining available layers. Notice that the layers with bad vertices can be chosen from all k + 1 layers except the $Rk - \alpha k + o(\beta)k$ layers which are chosen to have good vertices. Also, the inequality holds by application of the inequality $\binom{N}{\varepsilon N} < 2^{H(\varepsilon)N}$ for $0 < \varepsilon < \frac{1}{2}$ and positive integer N, where $H(\cdot)$ is the binary entropy function. Suppose that path Q has vertices from layers $\Lambda_{i_1}, \Lambda_{i_2}, \dots, \Lambda_{i_{Rk+o(\beta)k}}$. Let $\mathcal{I} := \{1, \dots, Rk + o(\beta)k\}$ be the set of indices of the layers with a vertex on the path Q. Let $\mathcal{I} = \mathcal{I}_g \cup \mathcal{I}_b$ where \mathcal{I}_g is the set of indices of layers with good vertices on Q and \mathcal{I}_b is the set of indices of layers with bad vertices on Q (The sets \mathcal{I}_g and \mathcal{I}_b are disjoint.). That is, layer Λ_{i_j} with $j \in \mathcal{I}_g$ is a layer with a bad vertex on Q.

Let us express the path Q as $s - v_{i_1} - v_{i_2} - \cdots - v_{i_{Rk+o(\beta)k}} - t$ where $v_{i_j} \in \Lambda_{i_j}$. Path Q is uniquely identified by the position of the first bit of its vertices, $(\check{v}_{i_1}, \check{v}_{i_2}, \cdots, \check{v}_{i_{Rk+o(\beta)k}})$. Equivalently, if we know the distance between consecutive vertices $(\text{Dis}(v_{i_j}, v_{i_{j+1}}) : j \in \mathcal{I})$, we can uniquely identify the position of each vertex on the path. Therefore, next we count the number of possible values of the distances between consecutive vertices $(\text{Dis}(v_{i_j}, v_{i_{j+1}}) : j \in \mathcal{I})$.

Since good vertices are pinned down on the path, the value of $\text{Dis}(v_{i_j}, v_{i_{j+1}})$ is determined if both v_{i_j} and $v_{i_{j+1}}$ are good vertices. Let us define set $\mathcal{H} \subset \mathcal{I}$ as follows

$$\mathcal{H} = \{ j : j \in \mathcal{I}_b \lor (j+1) \in \mathcal{I}_b \}$$

Therefore $(\text{Dis}(v_{i_j}, v_{i_{j+1}}) : j \in \mathcal{H})$ is the set of distances between consecutive vertices of Q that are undetermined. The number of bad vertices on Q is αk . Therefore $|\mathcal{H}| \leq 2\alpha k$. Let j_1 and j_2 with $j_1 < j_2$ be two consecutive elements in the ordered version of \mathcal{I}_g . Then by additivity of distances, bad vertices $v_{i_{j_1+1}}, \cdots, v_{i_{j_2-1}}$ need to satisfy the following constraint:

$$\sum_{t=j_1}^{j_2-1} \text{Dis}(v_{i_t}, v_{i_{t+1}}) = \text{Dis}(v_{i_{j_1}}, v_{i_{j_2}}) - (j_2 - j_1 - 1)L_P,$$
(8)

where $(j_2 - j_1 - 1)L_P$ is the total length of the substrings $v_{i_{j_1+1}}, \cdots, v_{i_{j_2-1}}$ in Y. Furthermore, bad vertices should be placed on Q such that they satisfy the constraint given in (6). For every $j \in \mathcal{H}$, we need to have

$$-1 \le \operatorname{Dis}(v_{i_j}, v_{i_{j+1}}) \le (i_{j+1} - i_j)L_S + (i_{j+1} - i_j - 1)L_P.$$
(9)

Next we find an upper bound on the number of integer vectors $(\text{Dis}(v_{i_j}, v_{i_{j+1}}) : j \in \mathcal{H})$ that satisfy (8) and (9).

For $j \in \mathcal{I}$, we use the following change of variables

$$\delta_j := (i_{j+1} - i_j)L_S + (i_{j+1} - i_j - 1)L_P - \operatorname{Dis}(v_{i_j}, v_{i_{j+1}}).$$

Equation (8) in terms of the variables δ_j 's is written as follows. For $j_1 < j_2$, as any two consecutive elements in the ordered version of \mathcal{I}_q , we have

$$\sum_{j=j_1}^{j_2-1} \delta_j = (i_{j_2} - i_{j_1})L_S + (i_{j_2} - i_{j_1} - 1)L_P - \operatorname{Dis}(v_{i_{j_1}}, v_{i_{j_2}})$$
(10)

Observe that in (10), $(i_{j_2} - i_{j_1})L_S + (i_{j_2} - i_{j_1} - 1)L_P$ is the number of bits between $P_{i_{j_1}}$ and $P_{i_{j_2}}$ in X and $\text{Dis}(v_{i_{j_1}}, v_{i_{j_2}})$ is the number of bits between the correct match of $P_{i_{j_1}}$ and the correct match of $P_{i_{j_2}}$ in Y. Therefore, the right hand side of Equation (10) is the number of deleted bits in the substring between $P_{i_{j_1}}$ and $P_{i_{j_2}}$ in X. To find an upper bound on the number of solutions for $(\delta_j : j \in \mathcal{H})$, we relax constraints

in (10) over all j's into a single constraint by adding them together:

$$\sum_{c \in \mathcal{H}} \delta_j = \delta - \sum_{j' \in \mathcal{H}^c} \delta_{j'}.$$
 (11)

Here δ is the total number of deleted bits from X and set $\mathcal{H}^c = \mathcal{I} \setminus \mathcal{H}$ is the set of indices j' for which $h_{j'}$ is determined; i.e., $v_{i_{j'}}$ and $v_{i_{j'+1}}$ are both good vertices. Furthermore, $\delta_{j'}$ is the number of deleted bits from the substring between $P_{i_{j'}}$ and $P_{i_{j'+1}}$ in X.

Next, we use the following result on the concentration of $\sum_{i \in \mathcal{H}} \delta_i$ around its expected value.

Lemma 8. For a random string X, random deletion pattern D, and the resultant string Y, the following bound holds:

$$\Pr\left\{\left|\sum_{j\in\mathcal{H}}\delta_j - \mathbb{E}\left[\sum_{j\in\mathcal{H}}\delta_j\right]\right| = o(\beta)k\right\} \ge 1 - 2^{-\Omega(n)}.$$

Proof: See Appendix II.

To estimate $\mathbb{E}\left[\sum_{j\in\mathcal{H}} \delta_j\right]$, first notice that the average number of deleted bits from X is $\mathbb{E}\left[\delta\right] = n\beta = (1 + o(\beta))k$.

Next we find $\mathbb{E}[\delta_{j'}]$ for $j' \in \mathcal{H}^c$. Since Q has $Rk + o(\beta)k$ vertices, the average size of the substring between $P_{i_{j'}}$ and $P_{i_{j'+1}}$ in X is $\frac{n}{Rk+o(\beta)k}$. Therefore, $\mathbb{E}[\delta_{j'}]$, the average number of deleted bits from the substring between $P_{i_{j'}}$ and $P_{i_{j'+1}}$ in X is

$$\mathbb{E}\left[\delta_{j'}\right] = \frac{n\beta}{Rk + o(\beta)k} = \frac{1}{R + o(\beta)} = 1 + \beta L_P - 2\beta + o(\beta).$$

Since $|\mathcal{H}| \leq 2\alpha k$ and $|\mathcal{I}| = |\mathcal{H}| + |\mathcal{H}^c| = (R + o(\beta))k$, we find that

$$|\mathcal{H}^c| \ge (R - 2\alpha + o(\beta))k = (1 - \beta L_P + 2\beta - 2\alpha + o(\beta))k.$$

We conclude that

$$\mathbb{E}[\sum_{j \in \mathcal{H}} \delta_j] = \mathbb{E}[\delta] - \mathbb{E}[\sum_{j' \in \mathcal{H}^c} \delta_{j'}]$$

= $k - |\mathcal{H}^c| \mathbb{E}[\delta_{j'}] + o(\beta)k$
 $\leq k \cdot$
 $(1 - (1 - \beta L_P + 2\beta - 2\alpha)(1 + \beta L_P - 2\beta) + o(\beta))$
= $2\alpha k (1 + \beta L_P - 2\beta) + o(\beta)k$,

and therefore by Lemma 8, with probability at least $1-2^{-\Omega(n)}$

$$\sum_{j \in \mathcal{H}} \delta_j = 2\alpha k (1 + \beta L_P - 2\beta) + o(\beta)k.$$
 (12)

Therefore, we showed that (8) yields the weaker constraint in (12) on the vector $(\delta_j : j \in \mathcal{H})$.

Now consider the inequality in (9). We can rewrite it in terms of δ_j as follows

$$0 \le \delta_j \le (i_{j+1} - i_j)L_S + (i_{j+1} - i_j - 1)L_P + 1.$$

To find an upper bound on the number of solutions for $(\delta_j : j \in \mathcal{H})$, we relax the preceding constraint to $\delta_j \ge 0$.

Under the constraint that $\delta_j \geq 0$, the number of integer solutions for $(\delta_j : j \in \mathcal{H})$ under the condition (12), is given by

$$\begin{pmatrix} 2\alpha k(1+\beta L_P-2\beta)+o(\beta)k+|\mathcal{H}|-1\\|\mathcal{H}|-1 \end{pmatrix} \leq \\ \begin{pmatrix} 2\alpha k(2+\beta L_P-2\beta+\frac{o(\beta)}{\alpha})\\2\alpha k \end{pmatrix} \leq \\ 2^{2\alpha k(2+\beta L_P-2\beta+\frac{o(\beta)}{\alpha})} \leq 2^{5\alpha k}, \quad (13)$$

where the last estimate holds for sufficiently small β .

Given the number of possibilities for path Q, we next compute the probability of occurrence of each realization of path Q. Since X is generated by an i.i.d. Bernoulli source of parameter $\frac{1}{2}$ and deletions occur independently, Y is also generated by an i.i.d. Bernoulli source of parameter $\frac{1}{2}$ and different substrings of Y are independent. Therefore, the probability of any given realization of bad vertices as specified by the choice of δ_j 's is $2^{-L_P\alpha k}$. By applying the union bound on the probability of existence of individual paths, and using inequalities (7) and (13), we conclude that the probability of the existence of a path Q with $Rk + o(\beta)k$ total vertices and αk bad vertices is upper bounded by $2^{\Delta_\alpha k}$ where

$$\begin{aligned} \Delta_{\alpha} &= RH\left(\frac{\alpha}{R}\right) + (1 - R + \alpha)H\left(\frac{\alpha}{1 - R + \alpha}\right) \\ &+ 5\alpha - L_{P}\alpha + o(\beta) \\ &= -\alpha\log\alpha + \alpha\log R - R\log(1 - \frac{\alpha}{R}) \\ &+ \alpha\log(1 - \frac{\alpha}{R}) - \alpha\log\alpha + \alpha\log(1 - R + \alpha) \\ &- (1 - R)\log(1 - \frac{\alpha}{1 - R + \alpha}) + 5\alpha - L_{P}\alpha + o(\beta) \end{aligned}$$

Next we find an upper bound for Δ_{α} . Since for sufficiently small β , R < 1, then $\alpha \log R < 0$. Since $\alpha < \frac{1}{2}$, for small enough β , $R > \alpha$. Therefore $\alpha \log(1 - \frac{\alpha}{R}) < 0$ and $\alpha \log(1 - R + \alpha) < 0$. Using the inequality $\log(1+x) \leq \frac{x}{\ln 2}$ for x > -1 we find that

$$-R\log(1-\frac{\alpha}{R}) = R\log(1+\frac{\alpha}{R-\alpha}) \le \frac{R\alpha}{(R-\alpha)\ln 2}$$
$$= \frac{\alpha}{(1-\frac{\alpha}{R})\ln 2}.$$

Notice that for small values of β , R is close to 1. Let us assume that β is sufficiently small such that R > 0.9. Since $\alpha < \frac{1}{2}$, we can write

$$-R\log(1-\frac{\alpha}{R}) \le \frac{\alpha}{(1-\frac{\alpha}{R})\ln 2} < \frac{\alpha}{(1-\frac{1}{1.8})\ln 2} = \frac{2.25\alpha}{\ln 2}.$$

Also we have

$$-(1-R)\log\left(1-\frac{\alpha}{1-R+\alpha}\right) = (1-R)\log\left(1+\frac{\alpha}{1-R}\right)$$
$$\leq \frac{(1-R)\alpha}{(1-R)\ln 2} = \frac{\alpha}{\ln 2}.$$

Therefore

$$\begin{aligned} \Delta_{\alpha} &\leq o(\beta) - 2\alpha \log \alpha + \frac{2.25\alpha}{\ln 2} + \frac{\alpha}{\ln 2} + 5\alpha - L_{P}\alpha \\ &= \alpha (\frac{o(\beta)}{\alpha} - 2\log \alpha + \frac{3.25}{\ln 2} + 5 - L_{P}) \\ &< \alpha (\frac{o(\beta)}{\alpha} - 2\log \alpha + 9.7 - L_{P}). \end{aligned}$$

Notice that $\frac{o(\beta)}{\alpha} \leq \frac{o(\beta)}{\beta}$ is arbitrarily small for sufficiently small β . If we choose β such that $\frac{o(\beta)}{\alpha} < 0.3$, then we have

$$\Delta_{\alpha} < \alpha(-2\log\alpha + 10 - L_P).$$

Case 2: $\frac{1}{2} \le \alpha \le R + o(\beta)$

We again seek to bound the probability of the existence of a path Q from s to t with $Rk + o(\beta)k$ total vertices and αk bad vertices. Let the path Q be denoted by $s - v_{i_1} - v_{i_2} - \cdots - v_{i_{Rk+o(\beta)k}} - t$ and let δ_j denote the number of deleted bits between vertices v_{i_j} and $v_{i_{j+1}}$. Clearly, the sum of δ_j is the total number of deletions in string Y. By Theorem 3, with probability at least $1 - 2^{-\Omega(n)}$ we have $\sum_{j=0}^{Rk+o(\beta)k} \delta_j = n\beta + n\beta o(\beta) = k(1 + o(\beta))$. The number of integer solutions for $\delta_j \ge 0$ under this constraint is

$$\binom{k+Rk+o(\beta)k-1}{Rk+o(\beta)k-1} \le \binom{2k}{k} \le 2^{2k}$$

The probability for each solution of δ_j 's to represent a valid s-t path is at most $2^{-L_P\alpha k}$. Therefore, an upper bound on the probability of existence of a path Q in this case is $2^{(2-L_P\alpha)k}$.

Finally, putting both cases for the range of α together, the probability of the existence of a path Q with $Rk + o(\beta)k$ vertices between s and t with at least βk bad vertices can be upper bounded by the sum of two integrals:

$$\int_{\alpha=\beta}^{\frac{1}{2}} 2^{\Delta_{\alpha}k} d\alpha k + \int_{\alpha=\frac{1}{2}}^{R+o(\beta)} 2^{(2-L_{P}\alpha)k} d\alpha k \leq \int_{\alpha=\beta}^{\frac{1}{2}} 2^{(-2\log\alpha+10-L_{P})\alpha k} d\alpha k + \int_{\alpha=\frac{1}{2}}^{R+o(\beta)} 2^{(2-L_{P}\alpha)k} d\alpha k.$$

If we pick $L_P \ge 11 + 2\log \frac{1}{\beta}$ then we find

$$(-2\log\alpha + 10 - L_P)\alpha k \le -\alpha k \le -\beta k$$

for $\beta \leq \alpha \leq \frac{1}{2}$. Also,

$$2 - L_P \alpha \le 2 - \frac{1}{2} \cdot 11 = -3.5.$$

Therefore, we can upper bound the sum of the two integrals by

$$\int_{\beta}^{\frac{1}{2}} k 2^{-\beta k} d\alpha + \int_{\frac{1}{2}}^{R+o(\beta)} k 2^{-3.5k} d\alpha \le \frac{k}{2} (2^{-\beta k} + 2^{-3.5k}) = 2^{-\Omega(n)}.$$

This completes the proof of Theorem 5.

In order to verify the result of Theorem 5 in a practical setting, we have plotted graph G for a randomly generated string X and a randomly generated deletion pattern D with parameter $\beta = 0.01$, for three values of L_P in Figure 4. To



Figure 4. Graph G for $k = 100, \beta = 0.01, L_S = 100$, and $L_P = 6, 7$, and 8, where only the edges between consecutive layers are depicted.

avoid visual complications, we have only plotted edges that connect vertices on two consecutive layers. As it is clear from the figure, for small values of L_P , there are many edges in the graph and there are potentially many paths that connect s to t which do not share many vertices with the correct path. However, for larger values of L_P , the irrelevant edges disappear from the graph and the only path that remains is the one formed by good vertices of the graph. For $\beta = 0.01$, Theorem 5 states that $L_P \geq 11 + 2\log \frac{1}{\beta} \approx 17$ is sufficient for our purpose. In practice, we observe values of L_P around 8 are sufficient for distinguishing good vertices on graph G.

IV. PRACTICAL IMPLEMENTATION

In this section we discuss practical implementation of our synchronization protocol, consisting of a matching module, a deletion recovery module, and an LDPC decoder module (see Figure 1). For the deletion recovery module, we can implement the synchronization protocol of Venkataramanan *et al.* [7] which runs in linear time in $|F_j|$ for deletion recovery of each substring F_j , $1 \le j \le k'$. Therefore the overall complexity of the deletion recovery module is linear in n. For the LDPC decoder module there are many sophisticated encoding and decoding schemes (see [15], [16]) that need running time linear in n.

In this section we therefore focus on the implementation of the graph-based algorithm for the matching module explained in the previous section. The result of Theorem 5 indicates that to find a large number of correct matches for pivots in the received string Y, it suffices to find an s - t path with $Rk + o(\beta)k$ vertices in the matching graph G. We now argue that this problem can be cast as the well known "shortest path problem" in a directed graph, so it can be efficiently solved in polynomial time.

As the first step, we only keep the vertices in graph G which have an edge to vertex t and remove all other vertices. Since all good vertices are connected to vertex t, this step does not eliminate any good vertex from graph G. Let \tilde{G} denote the resulting graph. As the second step, we find the longest s - tpath in \tilde{G} . Since all good vertices are connected together and form an s - t path of length $Rk + o(\beta)k$, the longest path in \tilde{G} has at least $Rk + o(\beta)k$ vertices. Finally, we modify the discovered path into a path with only $Rk + o(\beta)k$ vertices by keeping only the first $Rk + o(\beta)k$ vertices on the path. Since each vertex in graph \tilde{G} has an edge to vertex t, the resulting vertices from this step form a path with $Rk + o(\beta)k$ vertices from s to t.

The only step of the above procedure which is computationally demanding is the second step for finding the longest s-t path in \tilde{G} . Notice that since G and hence \tilde{G} are acyclic graphs, the longest s-t path problem in \tilde{G} , can be reduced to the shortest s-t path problem in \tilde{G} by assigning weight -1 to each edge. The latter problem is solvable in time $O(|\tilde{G}|^2)$, for instance by Dijkstra's algorithm [18], where $|\tilde{G}|$ is the number of vertices in \tilde{G} . We upper bound $|\tilde{G}|$ by |G|. To approximate |G|, we notice that there are $n\beta + o(1)$ layers in graph G and the number of vertices in layer Λ_i is the number of copies of pivot P_i in Y, which is on average $2^{-L_P}|Y| = O(\beta^2 n)$. Therefore, $|G| = O((\beta^2 n) \cdot (n\beta) + o(1)) = O(n^2\beta^3)$. We conclude that the complexity of matching pivots in graph Gis upper bounded by $O(|G|^2) = O(n^4\beta^6)$.

V. CONCLUSIONS

In this paper we offered the first synchronization protocol for recovering from a small rate of deletions with an optimal order of transmitted bits and with exponentially small reconstruction error. The main idea was to divide the synchronization problem into synchronization between shorter substrings of the source file and the destination file. For that, our protocol sends equally spaced small substrings of the source file to the destination, and destination then uses a graph theoretic algorithm to locate the short substrings within its file with high accuracy. For synchronization between the shorter substrings we used existing protocols that recover from a small number of edits. We observed that the compound output of the first two steps can be modeled as an output of a BSC with a small error probability. This error can be recovered with a low bit error rate by using an LDPC coding scheme.

While in this work we only considered recovering from i.i.d. patterns of deleted bits, there are many other interesting edit models that the ideas of this paper can be applied to. An immediate extension of our work is to the synchronization from i.i.d. insertions. To explain an i.i.d. insertion process, let us consider an equivalent description of the i.i.d. deletion process considered in this paper. In the new description, the deletion pattern D is described as an independent sequence of positive integers, where the integers alternatively represent the length of zero and one runs in the deletion pattern D. It is easy to verify that if the integers are generated independently according to an appropriate geometric distribution, the result is an i.i.d. 0-1 deletion pattern. We can describe the insertion pattern in the same way by generating the run length sequence of the pattern. For the insertion pattern, each run of ones corresponds to an inserted substring of equal length generated by an i.i.d. Bernoulli process. Also, each run of zeros corresponds to a substring of the input string of equal length in the output. It is not hard to see that the solution presented in this paper for synchronization from deletions is directly applicable to solving the synchronization problem from random insertions. One can also consider more general patterns of deletions or insertions, e.g., the 0-1 deletion (insertion) patterns that follow a Markov chain random process (see [11]).

Another interesting direction for the extension of this work is the design of synchronization protocols that are capable of recovering from a small rate of both deletions and insertions. While the deletion recovery module in our work, based on the algorithm by Venkataramanan *et al.* [7], is directly applicable to recovery from deletions and insertions, the main challenge is to extend the graph theoretic algorithm for matching the pivot substrings in the received string Y when there are both deletions and insertions. Again, many parts of our argument still hold for the new setting as long as the edits happen with small rates while some technical parts may need to be modified. This extension is the focus of our current research. Our recent progress is reported in [19].

There are some other aspects of our current research that can be modified into more efficient synchronization protocols. For example, our algorithm needs a small backward bandwidth from node B to node A in the deletion recovery module. This bandwidth is an inherent component of the synchronization protocol of Venkataramanan *et al.*, [7]. It is of great interest to design protocols that can operate on forward links only. As proved by Orlitsky [3], design of optimal protocols for recovery from deletions on forward links implies optimal protocols for recovery from deletions and insertions. Furthermore, such protocols can be implemented as efficient channel codes for communicating over edit channels (see [11], [20], [21], [22]).

Finally, from a practical perspective, it is interesting to design a more efficient implementation of the graph theoretic matching algorithm which is at the heart of our matching module. While our algorithm runs in $O(n^4\beta^6)$ time, we believe that by exploiting the specific structure of the matching graph, and applying additional restrictions on the connectivity of the vertices of the graph together, it is possible to considerably reduce the running time of the matching module and hence reduce the overall complexity of the synchronization protocol.

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APPENDIX I

Here we evaluate $\mathbb{E} \left[\delta_i \log |F_i| \right]$.

$$\mathbb{E}\left[\delta_{j} \log |F_{j}|\right] = \sum_{l} \Pr\left\{|F_{j}| = l\right\} \mathbb{E}\left[\delta_{j} \log |F_{j}| \left| |F_{j}| = l\right]$$
$$= \sum_{l} \beta l \log l \Pr\left\{|F_{j}| = l\right\}$$
$$= \mathbb{E}\left[\beta|F_{j}| \log|F_{j}|\right].$$

Next we estimate $\mathbb{E}\left[\beta|F_j|\log|F_j|\right]$. Recall that F_j is the substring of X between $P_{i_{j-1}}$ and P_{i_j} . There are $(i_j - i_{j-1})$ segment strings and $(i_j - i_{j-1} - 1)$ pivot strings between $P_{i_{j-1}}$ and P_{i_i} . Therefore

$$|F_j| = (i_j - i_{j-1})L_S + (i_j - i_{j-1} - 1)L_P.$$

There is a total of k pivots, and k' of them are matched by the matching module. Therefore, with probability $p := \frac{k'}{k} =$ $(1 - L_P\beta + 2\beta + o(\beta))$ pivot P_i is matched. Furthermore, the probability that a pivot is matched is independent of other pivots. Thus, $(i_j - i_{j-1})$ has the following geometric distribution

$$\Pr\{i_j - i_{j-1} = \ell\} = p(1-p)^{\ell-1}$$

Suppose $r \in \{1, 2, \dots\}$ is a random variable distributed as above. If we upper bound $|F_j| \leq (i_j - i_{j-1})(L_S + L_P)$, then

$$\mathbb{E}\left[\beta|F_j|\log|F_j|\right] \le \mathbb{E}\left[\beta r(L_S + L_P)\log r(L_S + L_P)\right]$$

= $\beta(L_S + L_P)\mathbb{E}\left[r\log r + r\log(L_S + L_P)\right]$
 $\le 2\mathbb{E}\left[r^2\right] + 2\log(L_S + L_P)\mathbb{E}\left[r\right],$

where we used the fact that $\beta(L_S + L_P) \leq 2$ and $r \log r \leq r^2$. We can write

$$\mathbb{E}\left[r\right] = \frac{1}{p}, \mathbb{E}\left[r^{2}\right] = \operatorname{Var}(r) + \mathbb{E}\left[r\right]^{2} = \frac{2-p}{p^{2}}$$

Also, we use $\log(L_S + L_P) \leq \log 2L_S \leq 2 \log \frac{1}{\beta}$ and find that

$$\mathbb{E}\left[\beta|F_{j}|\log|F_{j}|\right] \le \frac{4-2p}{p^{2}} + \frac{4}{p}\log\frac{1}{\beta} \le 16 + 8\log\frac{1}{\beta},$$

where we used the fact that $\frac{4-2p}{p^2} \le 16$ and $\frac{4}{p} \le 8$ for $p \ge \frac{1}{2}$ (Notice that $p \to 1$, as $\beta \to 0$.).

APPENDIX II

Recall that δ_j is the number of deleted bits from the substring of X between pivots P_{i_j} and $P_{i_{j+1}}$. Let us denote by \mathcal{L}_P the set of indices l for which P_l appears between P_{i_i} and $P_{i_{j+1}}$ for some $j \in \mathcal{H}$. Similarly, let \mathcal{L}_S denote the set of indices l for which S_l appears between P_{i_j} and $P_{i_{j+1}}$ for some $j \in \mathcal{H}$. Let δ_{P_l} denote the number of deleted bits from P_l and δ_{S_l} denote the number of deleted bits from S_l . We can write

$$\sum_{j \in \mathcal{H}} \delta_j = \sum_{l \in \mathcal{L}_P} \delta_{P_l} + \sum_{l \in \mathcal{L}_S} \delta_{S_l}$$

Notice that the length of the interval that δ_{P_l} takes values from is $L_P = O(\log \frac{1}{\beta})$ and the length of the interval that δ_{S_l} takes values from is $L_S = \frac{1}{\beta}$. Next, by application of Theorem 4 we can write

$$\begin{split} &\Pr\left\{\left|\sum_{j\in\mathcal{H}}\delta_{j}-\mathbb{E}\left[\sum_{j\in\mathcal{H}}\delta_{j}\right]\right|=o(\beta)k\right\}\geq\\ &\Pr\left\{\left|\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}\right]\right|+\\ &\left|\sum_{l\in\mathcal{L}_{S}}\delta_{S_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}\right]\right|=o(\beta)k\right\}=\\ &\Pr\left\{\left|\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}\right]\right|=o(\beta)k\right\}\cdot\\ &\Pr\left\{\left|\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}\right]\right|=o(\beta)k\right\}-\\ &\Pr\left\{\left|\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{P}}\delta_{P_{l}}\right]\right|=\frac{o(\beta)k}{|\mathcal{L}_{P}|}\right\}\cdot\\ &\Pr\left\{\left|\sum_{l\in\mathcal{L}_{S}}\delta_{S_{l}}-\mathbb{E}\left[\sum_{l\in\mathcal{L}_{S}}\delta_{S_{l}}\right]\right|=\frac{o(\beta)k}{|\mathcal{L}_{S}|}|\mathcal{L}_{S}|\right\}\geq\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}|\mathcal{L}_{P}|}{|\mathcal{L}_{P}|^{2}L_{P}^{2}})\right)\cdot\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}|\mathcal{L}_{S}|}{|\mathcal{L}_{P}|\mathcal{L}_{P}^{2}})\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k}{|\mathcal{L}_{P}|\mathcal{L}_{P}^{2}})\right)\cdot\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}}{|\mathcal{L}_{P}|\mathcal{L}_{P}^{2}})\right)\geq\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}}{\mathcal{L}_{S}^{2}})\right)\geq\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}}{\mathcal{L}_{S}^{2}})\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}}{\mathcal{L}_{S}^{2}})\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})k}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-\frac{2o(\beta^{2})k^{2}}{\mathcal{L}_{S}^{2}})\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})\beta}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-\frac{2o(\beta^{2})k}{\mathcal{L}_{S}^{2}})\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})\beta}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-2\beta^{3}o(\beta^{2})n)\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})\beta}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-2\beta^{3}o(\beta^{2})n)\right)=\\ &\left(1-2\exp(-\frac{2o(\beta^{2})\beta}{\mathcal{L}_{P}^{2}})\right)\cdot\left(1-2\exp(-2\beta^{3}o(\beta^{2})n)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2}))\right)\cdot\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)\cdot\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)\cdot\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)=\\ &\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)+\left(1-2\exp(-2\beta^{2}n(\beta^{2})n\right)\right)$$

where in our derivation we used the fact that $|\mathcal{L}_P| \leq k$ and

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